INTRODUCTION TO THEORETICAL COMPUTER SCIENCE

TEXTBOOK IN PREPARATION.
AVAILABLE ON HTTPS://INTROTCS.ORG
To Ravit, Alma and Goren.
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Preface

“We make ourselves no promises, but we cherish the hope that the unobstructed pursuit of useless knowledge will prove to have consequences in the future as in the past” ...

“An institution which sets free successive generations of human souls is amply justified whether or not this graduate or that makes a so-called useful contribution to human knowledge. A poem, a symphony, a painting, a mathematical truth, a new scientific fact, all bear in themselves all the justification that universities, colleges, and institutes of research need or require”, Abraham Flexner, The Usefulness of Useless Knowledge, 1939.

“I suggest that you take the hardest courses that you can, because you learn the most when you challenge yourself... CS 121 I found pretty hard.”, Mark Zuckerberg, 2005.

This is a textbook for an undergraduate introductory course on Theoretical Computer Science. The educational goals of this book are to convey the following:

- That computation arises in a variety of natural and human-made systems, and not only in modern silicon-based computers.

- Similarly, beyond being an extremely important tool, computation also serves as a useful lens to describe natural, physical, mathematical and even social concepts.

- The notion of universality of many different computational models, and the related notion of the duality between code and data.

- The idea that one can precisely define a mathematical model of computation, and then use that to prove (or sometimes only conjecture) lower bounds and impossibility results.

- Some of the surprising results and discoveries in modern theoretical computer science, including the prevalence of NP-completeness, the power of interaction, the power of randomness on one hand
and the possibility of derandomization on the other, the ability to use hardness “for good” in cryptography, and the fascinating possibility of quantum computing.

I hope that following this course, students would be able to recognize computation, with both its power and pitfalls, as it arises in various settings, including seemingly “static” content or “restricted” formalisms such as macros and scripts. They should be able to follow through the logic of proofs about computation, including the central concept of a reduction, as well as understanding “self-referential” proofs (such as diagonalization-based proofs that involve programs given their own code as input). Students should understand that some problems are inherently intractable, and be able to recognize the potential for intractability when they are faced with a new problem. While this book only touches on cryptography, students should understand the basic idea of how we can use computational hardness for cryptographic purposes. However, more than any specific skill, this book aims to introduce students to a new way of thinking of computation as an object in its own right and to illustrate how this new way of thinking leads to far-reaching insights and applications.

My aim in writing this text is to try to convey these concepts in the simplest possible way and try to make sure that the formal notation and model help elucidate, rather than obscure, the main ideas. I also tried to take advantage of modern students’ familiarity (or at least interest!) in programming, and hence use (highly simplified) programming languages to describe our models of computation. That said, this book does not assume fluency with any particular programming language, but rather only some familiarity with the general notion of programming. We will use programming metaphors and idioms, occasionally mentioning specific programming languages such as Python, C, or Lisp, but students should be able to follow these descriptions even if they are not familiar with these languages.

Proofs in this book, including the existence of a universal Turing Machine, the fact that every finite function can be computed by some circuit, the Cook-Levin theorem, and many others, are often constructive and algorithmic, in the sense that they ultimately involve transforming one program to another. While it is possible to follow these proofs without seeing the code, I do think that having access to the code, and the ability to play around with it and see how it acts on various programs, can make these theorems more concrete for the students. To that end, an accompanying website (which is still work in progress) allows executing programs in the various computational models we define, as well as see constructive proofs of some of the theorems.
0.1 TO THE STUDENT

This book can be challenging, mainly because it brings together a variety of ideas and techniques in the study of computation. There are quite a few technical hurdles to master, whether it is following the diagonalization argument in proving the Halting Problem is undecidable, combinatorial gadgets in NP-completeness reductions, analyzing probabilistic algorithms, or arguing about the adversary to prove the security of cryptographic primitives.

The best way to engage with this material is to read these notes actively, so make sure you have a pen ready. While reading, I encourage you to stop and think about the following:

• When I state a theorem, stop and take a shot at proving it on your own before reading the proof. You will be amazed by how much you can understand a proof better even after only 5 minutes of attempting it on your own.

• When reading a definition, make sure that you understand what the definition means, and what are natural examples of objects that satisfy it and objects that do not. Try to think of the motivation behind the definition, and whether there are other natural ways to formalize the same concept.

• Actively notice which questions arise in your mind as you read the text, and whether or not they are answered in the text.

As a general rule, it is more important that you understand the definitions than the theorems, and it is more important that you understand a theorem statement than its proof. After all, before you can prove a theorem, you need to understand what it states, and to understand what a theorem is about, you need to know the definitions of the objects involved. Whenever a proof of a theorem is at least somewhat complicated, I provide a “proof idea.” Feel free to skip the actual proof in a first reading, focusing only on the proof idea.

This book contains some code snippets, but this is by no means a programming text. You don’t need to know how to program to follow this material. The reason we use code is that it is a precise way to describe computation. Particular implementation details are not as important to us, and so we will emphasize code readability at the expense of considerations such as error handling, encapsulation, etc. that can be extremely important for real-world programming.

0.1.1 Is the effort worth it?

This is not an easy book, and you might reasonably wonder why should you spend the effort in learning this material. A traditional
justification for a “Theory of Computation” course is that you might encounter these concepts later on in your career. Perhaps you will come across a hard problem and realize it is NP complete, or find a need to use what you learned about regular expressions. This might very well be true, but the main benefit of this book is not in teaching you any practical tool or technique, but instead in giving you a different way of thinking: an ability to recognize computational phenomena even when they occur in non-obvious settings, a way to model computational tasks and questions, and to reason about them.

Regardless of any use you will derive from this book. I believe learning this material is important because it contains concepts that are both beautiful and fundamental. The role that energy and matter played in the 20th century is played in the 21st by computation and information, not just as tools for our technology and economy, but also as the basic building blocks we use to understand the world. This book will give you a taste of some of the theory behind those, and hopefully spark your curiosity to study more.

0.2 TO POTENTIAL INSTRUCTORS

I wrote this book for my Harvard course, but I hope that other lecturers will find it useful as well. To some extent, it is similar in content to “Theory of Computation” or “Great Ideas” courses such as those taught at CMU or MIT.

The most significant difference between our approach and more traditional ones (such as Hopcroft and Ullman’s [HU69; HU79] and Sipser’s [Sip97]) is that we do not start with finite automata as our initial computational model. Instead, our initial computational model is Boolean Circuits.¹ We believe that Boolean Circuits are more fundamental to the theory of computing (and even its practice!) than automata. In particular, Boolean Circuits are a prerequisite for many concepts that one would want to teach in a modern course on Theoretical Computer Science, including cryptography, quantum computing, derandomization, attempts at proving $P \neq NP$, and more. Even in cases where Boolean Circuits are not strictly required, they can often offer significant simplifications (as in the case of the proof of the Cook-Levin Theorem).

Furthermore, I believe there are pedagogical reasons to start with Boolean circuits as opposed to finite automata. Boolean circuits are a more natural model of computation, and one that corresponds more closely to computing in Silicon, making the connection to practice more immediate to the students. Finite functions are arguably easier to grasp than infinite ones, as we can fully write down their truth table. The theorem that every finite function can be computed by some Boolean circuit is both simple enough and important enough to serve

¹ An earlier book that starts with circuits as the initial model is John Savage’s [Sav98].
as an excellent starting point for this course. Moreover, many of the 
main conceptual points of the theory of computation, including the 
notions of the duality between code and data, and the idea of universal-
ity, can already be seen in this context.

After Boolean circuits, we move on to Turing machines and prove 
results such as the existence of a universal Turing machine, the un-
computability of the halting problem, and Rice’s Theorem. Automata 
are discussed after we see Turing machines and undecidability, as an 
example for a restricted computational model where problems such as 
determining halting can be effectively solved.

While this is not our motivation, the order we present circuits, Tur-
ing machines, and automata roughly corresponds to the chronological 
order of their discovery. Boolean algebra goes back to Boole’s and 
DeMorgan’s works in the 1840s [Boo47; De 47] (though the definition 
of Boolean circuits and the connection to physical computation was 
given 90 years later by Shannon [Shannon1937]). Alan Turing defined 
what we now call “Turing Machines” in the 1930s [Tur37], while finite 
automata were introduced in the 1943 work of McCulloch and Pitts 
[MP43] but only really understood in the seminal 1959 work of Rabin 
and Scott [RS59].

More importantly, while models such as finite-state machines, reg-
ular expressions, and context-free grammars are incredibly important 
for practice, the main applications for these models (whether it is for 
parsing, for analyzing properties such as liveness and safety, or even for 
software defined routing tables) rely crucially on the fact that these 
are tractable models for which we can effectively answer semantic ques-
tions. This practical motivation can be better appreciated after students 
see the undecidability of semantic properties of general computing 
models.

The fact that we start with circuits makes proving the Cook-Levin 
Theorem much easier. In fact, our proof of this theorem can (and is) 
done using a handful of lines of Python. Combining this proof with 
the standard reductions (which are also implemented in Python) 
allows students to appreciate visually how a question about computa-
tion can be mapped into a question about (for example) the existence 
of an independent set in a graph.

Some other differences between this book and previous texts are 
the following:

1. For measuring time complexity, we use the standard RAM machine 
model used (implicitly) in algorithms courses, rather than Tur-
ing machines. While these two models are of course polynomially 
equivalent, and hence make no difference for the definitions of the 
classes P, NP, and EXP, our choice makes the distinction between
notions such as $O(n)$ or $O(n^2)$ time more meaningful. This choice also ensures that these finer-grained time complexity classes correspond to the informal definitions of linear and quadratic time that students encounter in their algorithms lectures (or their whiteboard coding interviews.).

2. We use the terminology of functions rather than languages. That is, rather than saying that a Turing Machine $M$ decides a language $L \subseteq \{0, 1\}^*$, we say that it computes a function $F : \{0, 1\}^* \rightarrow \{0, 1\}$. The terminology of “languages” arises from Chomsky’s work [Cho56], but it is often more confusing than illuminating. The language terminology also makes it cumbersome to discuss concepts such as algorithms that compute functions with more than one bit of output (including basic tasks such as addition, multiplication, etc.). The fact that we use functions rather than languages means we have to be extra vigilant about students distinguishing between the specification of a computational task (e.g., the function) and its implementation (e.g., the program). On the other hand, this point is so important that it is worth repeatedly emphasizing and drilling into the students, regardless of the notation used. The book does mention the language terminology and reminds of it occasionally, to make it easier for students to consult outside resources.

Reducing the time dedicated to finite automata and context-free languages allows instructors to spend more time on topics that a modern course in the theory of computing needs to touch upon. These include randomness and computation, the interactions between proofs and programs (including Gödel’s incompleteness theorem, interactive proof systems, and even a bit on the λ-calculus and the Curry-Howard correspondence), cryptography, and quantum computing.

This book contains sufficient detail to enable its use for self-study. Toward that end, every chapter starts with a list of learning objectives, ends with a recap, and is peppered with “pause boxes” which encourage students to stop and work out an argument or make sure they understand a definition before continuing further.

Section 0.5 contains a “roadmap” for this book, with descriptions of the different chapters, as well as the dependency structure between them. This can help in planning a course based on this book.

0.3 ACKNOWLEDGEMENTS

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If you have any comments or suggestions, please do post them on the GitHub repository https://github.com/boazbk/tcs.

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Finally, I would like to thank my family: my wife Ravit, and my children Alma and Goren. Working on this book (and the corresponding course) took so much of my time that Alma wrote an essay for her fifth-grade class saying that “universities should not pressure professors to work too much.” I’m afraid all I have to show for this effort is 600 pages of ultra-boring mathematical text.
PRELIMINARIES
0

Introduction

“Computer Science is no more about computers than astronomy is about telescopes”, attributed to Edsger Dijkstra. ¹

“Hackers need to understand the theory of computation about as much as painters need to understand paint chemistry.”, Paul Graham 2003. ²

“The subject of my talk is perhaps most directly indicated by simply asking two questions: first, is it harder to multiply than to add? and second, why?...I (would like to) show that there is no algorithm for multiplication computationally as simple as that for addition, and this proves something of a stumbling block.”, Alan Cobham, 1964

The origin of much of science and medicine can be traced back to the ancient Babylonians. However, the Babylonians’ most significant contribution to humanity was arguably the invention of the place-value number system. The place-value system represents any number using a collection of digits, whereby the position of the digit is used to determine its value, as opposed to a system such as Roman numerals, where every symbol has a fixed numerical value regardless of position. For example, the average distance to the moon is approximately 238,900 of our miles or 259,956 Roman miles. The latter quantity, expressed in standard Roman numerals is

```
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
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MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
MMMMMMMMMMMMMMMMMMMMMMMMMMMMMM
``` ¹

1 This quote is typically read as disparaging the importance of actual physical computers in Computer Science, but note that telescopes are absolutely essential to astronomy as they provide us with the means to connect theoretical predictions with actual experimental observations.

2 To be fair, in the following sentence Graham says “you need to know how to calculate time and space complexity and about Turing completeness”. This book includes these topics, as well as others such as NP-hardness, randomization, cryptography, quantum computing, and more.

Learning Objectives:

- Introduce and motivate the study of computation for its own sake, irrespective of particular implementations.
- The notion of an algorithm and some of its history.
- Algorithms as not just tools, but also ways of thinking and understanding.
- Taste of Big-Ω analysis and the surprising creativity in the design of efficient algorithms.
Writing the distance to the sun in Roman numerals would require about 100,000 symbols: a 50-page book just containing this single number!

For someone who thinks of numbers in an additive system like Roman numerals, quantities like the distance to the moon or sun are not merely large—they are *unspeakable*: cannot be expressed or even grasped. It’s no wonder that Eratosthenes, who was the first person to calculate the earth’s diameter (up to about ten percent error) and Hipparchus who was the first to calculate the distance to the moon, did not use a Roman-numeral type system but rather the Babylonian sexadecimal (i.e., base 60) place-value system.

### 0.1 INTEGER MULTIPLICATION: AN EXAMPLE OF AN ALGORITHM

In the language of Computer Science, the place-value system for representing numbers is known as a *data structure*: a set of instructions or “recipe” for representing objects as symbols. An *algorithm* is a set of instructions or recipe for performing operations on such representations. Data structures and algorithms have enabled amazing applications that have transformed human society, but their importance goes beyond their practical utility. Structures from computer science, such as bits, strings, graphs, and even the notion of a program itself, as well as concepts such as universality and replication, have not just found (many) practical uses but contributed a new language and a new way to view the world.

In addition to coming up with the place-value system, the Babylonians also invented the “standard algorithms” that we were all taught in elementary school for adding and multiplying numbers. These algorithms have been essential throughout the ages for people using abaci, papyrus, or pencil and paper, but in our computer age, do they still serve any purpose beyond torturing third graders? To see why these algorithms are still very much relevant, let us compare the Babylonian digit-by-digit multiplication algorithm (“grade-school multiplication”) with the naive algorithm that multiplies numbers through repeated addition. We start by formally describing both algorithms, see Algorithm 0.1 and Algorithm 0.2.
Algorithm 0.1 — Multiplication via repeated addition.

Input: Non-negative integers $x, y$
Output: Product $x \cdot y$
1: Let $\text{result} \leftarrow 0$.
2: for $i = 1, \ldots, y$ do
3: $\text{result} \leftarrow \text{result} + x$
4: end for
5: return $\text{result}$

Algorithm 0.2 — Grade-school multiplication.

Input: Non-negative integers $x, y$
Output: Product $x \cdot y$
1: Write $x = x_{n-1}x_{n-2}\cdots x_0$ and $y = y_{m-1}y_{m-2}\cdots y_0$ in decimal place-value notation. \# $x_0$ is the ones digit of $x$, $x_1$ is the tens digit, etc.
2: Let $\text{result} \leftarrow 0$
3: for $i = 0, \ldots, n - 1$ do
4: for $j = 0, \ldots, m - 1$ do
5: $\text{result} \leftarrow \text{result} + 10^{i+j} \cdot x_i \cdot y_j$
6: end for
7: end for
8: return $\text{result}$

Both Algorithm 0.1 and Algorithm 0.2 assume that we already know how to add numbers, and Algorithm 0.2 also assumes that we can multiply a number by a power of 10 (which is, after all, a simple shift). Suppose that $x$ and $y$ are two integers of $n = 20$ decimal digits each. (This roughly corresponds to 64 binary digits, which is a common size in many programming languages.) Computing $x \cdot y$ using Algorithm 0.1 entails adding $x$ to itself $y$ times which entails (since $y$ is a 20-digit number) at least $10^{19}$ additions. In contrast, the grade-school algorithm (i.e., Algorithm 0.2) involves $n^2$ shifts and single-digit products, and so at most $2n^2 = 800$ single-digit operations. To understand the difference, consider that a grade-schooler can perform a single-digit operation in about 2 seconds, and so would require about 1,600 seconds (about half an hour) to compute $x \cdot y$ using Algorithm 0.2. In contrast, even though it is more than a billion times faster than a human, if we used Algorithm 0.1 to compute $x \cdot y$ using a modern PC, it would take us $10^{20}/10^9 = 10^{11}$ seconds (which is more than three millennia!) to compute the same result.

Computers have not made algorithms obsolete. On the contrary, the vast increase in our ability to measure, store, and communicate
data has led to much higher demand for developing better and more sophisticated algorithms that empower us to make better decisions based on these data. We also see that in no small extent the notion of algorithm is independent of the actual computing device that executes it. The digit-by-digit multiplication algorithm is vastly better than iterated addition, regardless whether the technology we use to implement it is a silicon-based chip, or a third grader with pen and paper.

Theoretical computer science is concerned with the inherent properties of algorithms and computation; namely, those properties that are independent of current technology. We ask some questions that were already pondered by the Babylonians, such as “what is the best way to multiply two numbers?”, but also questions that rely on cutting-edge science such as “could we use the effects of quantum entanglement to factor numbers faster?”.

**Remark 0.3 — Specification, implementation and analysis of algorithms.** A full description of an algorithm has three components:

- **Specification**: What is the task that the algorithm performs (e.g., multiplication in the case of Algorithm 0.1 and Algorithm 0.2.)
- **Implementation**: How is the task accomplished: what is the sequence of instructions to be performed. Even though Algorithm 0.1 and Algorithm 0.2 perform the same computational task (i.e., they have the same specification), they do it in different ways (i.e., they have different implementations).
- **Analysis**: Why does this sequence of instructions achieve the desired task. A full description of Algorithm 0.1 and Algorithm 0.2 will include a proof for each one of these algorithms that on input $x, y$, the algorithm does indeed output $x \cdot y$.

Often as part of the analysis we show that the algorithm is not only correct but also efficient. That is, we want to show that not only will the algorithm compute the desired task, but will do so in prescribed number of operations. For example Algorithm 0.2 computes the multiplication function on inputs of $n$ digits using $O(n^2)$ operations, while Algorithm 0.4 (described below) computes the same function using $O(n^{1.6})$ operations. (We define the $O$ notations used here in Section 1.4.8.)
0.2 Extended Example: A Faster Way to Multiply (Optional)

Once you think of the standard digit-by-digit multiplication algorithm, it seems like the “obviously best” way to multiply numbers. In 1960, the famous mathematician Andrey Kolmogorov organized a seminar at Moscow State University in which he conjectured that every algorithm for multiplying two \( n \) digit numbers would require a number of basic operations that is proportional to \( n^2 \) (\( \Omega(n^2) \) operations, using \( O \)-notation as defined in Chapter 1). In other words, Kolmogorov conjectured that in any multiplication algorithm, doubling the number of digits would quadruple the number of basic operations required. A young student named Anatoly Karatsuba was in the audience, and within a week he disproved Kolmogorov’s conjecture by discovering an algorithm that requires only about \( Cn^{1.6} \) operations for some constant \( C \). Such a number becomes much smaller than \( n^2 \) as \( n \) grows and so for large \( n \) Karatsuba’s algorithm is superior to the grade-school one. (For example, Python’s implementation switches from the grade-school algorithm to Karatsuba’s algorithm for numbers that are 1000 bits or larger.) While the difference between an \( O(n^{1.6}) \) and an \( O(n^2) \) algorithm can be sometimes crucial in practice (see Section 0.3 below), in this book we will mostly ignore such distinctions. However, we describe Karatsuba’s algorithm below since it is a good example of how algorithms can often be surprising, as well as a demonstration of the analysis of algorithms, which is central to this book and to theoretical computer science at large.

Karatsuba’s algorithm is based on a faster way to multiply two-digit numbers. Suppose that \( x, y \in [100] = \{0, \ldots , 99\} \) are a pair of two-digit numbers. Let’s write \( \bar{x} \) for the “tens” digit of \( x \), and \( \bar{x} \) for the “ones” digit, so that \( x = 10\bar{x} + x \), and write similarly \( y = 10\bar{y} + y \) for \( \bar{y}, y \in [10] \). The grade-school algorithm for multiplying \( x \) and \( y \) is illustrated in Fig. 1.

Generally, in the grade-school algorithm doubling the number of digits in the input results in quadrupling the number of operations, leading to an \( O(n^2) \) times algorithm. In contrast, Karatsuba’s algorithm is based on the observation that we can express Eq. (1) also as

\[
(10\bar{x} + x) \times (10\bar{y} + y) = (100\bar{x}\bar{y} + 10(\bar{x}y + x\bar{y})) + xy
\]  

(1)

Figure 1: The grade-school multiplication algorithm illustrated for multiplying \( x = 10\bar{x} + x \) and \( y = 10\bar{y} + y \). It uses the formula \( (10\bar{x} + x) \times (10\bar{y} + y) = 100\bar{x}\bar{y} + 10(x\bar{y} + y\bar{x}) + xy \).
which reduces multiplying the two-digit number $x$ and $y$ to computing the following three simpler products: $xy$, $x\overline{y}$ and $(\overline{x} + x)(\overline{y} + y)$.

By repeating the same strategy recursively, we can reduce the task of multiplying two $n$-digit numbers to the task of multiplying three pairs of $\lfloor n/2 \rfloor + 1$ digit numbers. Since every time we double the number of digits we triple the number of operations, we will be able to multiply numbers of $n = 2^\ell$ digits using about $3^\ell = n \log_2 3 \approx n^{1.585}$ operations.

The above is the intuitive idea behind Karatsuba’s algorithm, but is not enough to fully specify it. A complete description of an algorithm entails a precise specification of its operations together with its analysis: proof that the algorithm does in fact do what it’s supposed to do. The operations of Karatsuba’s algorithm are detailed in Algorithm 0.4, while the analysis is given in Lemma 0.5 and Lemma 0.6.

Algorithm 0.4 — Karatsuba multiplication.

Input: nonnegative integers $x, y$ each of at most $n$ digits

Output: $x \cdot y$

1: procedure Karatsuba($x, y$)
2: if $n \leq 2$ then
3: return $x \cdot y$
4: end if
5: Let $m = \lfloor n/2 \rfloor$
6: Write $x = 10^m \overline{x} + \underline{x}$ and $y = 10^m \overline{y} + \underline{y}$
7: $A \leftarrow$ Karatsuba($\overline{x}, \overline{y}$)
8: $B \leftarrow$ Karatsuba($\overline{x} + \underline{x}, \overline{y} + \underline{y}$)
9: $C \leftarrow$ Karatsuba($\underline{x}, \underline{y}$)
10: return $(10^n - 10^m) \cdot A + 10^m \cdot B + (1 - 10^m) \cdot C$
11: end procedure

Algorithm 0.4 is only half of the full description of Karatsuba’s algorithm. The other half is the analysis, which entails proving that (1) Algorithm 0.4 indeed computes the multiplication operation and (2) it does so using $O(n \log_2 3)$ operations. We now turn to showing both facts:

Lemma 0.5 For every nonnegative integers $x, y$, when given input $x, y$ Algorithm 0.4 will output $x \cdot y$.

Proof. Let $n$ be the maximum number of digits of $x$ and $y$. We prove the lemma by induction on $n$. The base case is $n \leq 2$ where the algorithm returns $x \cdot y$ by definition. Otherwise, if $n > 2$, we define $m = \lfloor n/2 \rfloor$, and write $x = 10^m \overline{x} + \underline{x}$ and $y = 10^m \overline{y} + \underline{y}$.

Plugging this into $x \cdot y$, we get

$$x \cdot y = 10^{2m} \overline{xy} + 10^m (\overline{x}y + x\overline{y}) + xy.$$  

If $x$ is a number then $\lfloor x \rfloor$ is the integer obtained by rounding it down, see Section 1.7.
Rearranging the terms we see that

\[ x \cdot y = 10^{2m} \overline{xy} + 10^m \left[ (\overline{x} + \overline{y})(\overline{y} + y) - xy - \overline{xy} \right] + xy. \quad (4) \]

since the numbers \( \overline{x}, \overline{y}, \overline{xy}, \overline{y} + y \) all have at most \( m + 1 < n \) digits, the induction hypothesis implies that the values \( A, B, C \) computed by the recursive calls will satisfy \( A = \overline{xy}, B = (\overline{x} + \overline{y})(\overline{y} + y) \) and \( C = xy \). Plugging this into (4) we see that \( x \cdot y \) equals the value \((10^{2m} - 10^m) \cdot A + 10^m \cdot B + (1 - 10^m) \cdot C \) computed by Algorithm 0.4.

\[ \text{Lemma 0.6} \] If \( x, y \) are integers of at most \( n \) digits, Algorithm 0.4 will take \( O(n \log_3^3) \) operations on input \( x, y \).

\[ \text{Proof.} \] Fig. 2 illustrates the idea behind the proof, which we only sketch here, leaving filling out the details as Exercise 0.4. The proof is again by induction. We define \( T(n) \) to be the maximum number of steps that Algorithm 0.4 takes on inputs of length at most \( n \). Since in the base case \( n \leq 2 \), Exercise 0.4 performs a constant number of computation, we know that \( T(2) \leq c \) for some constant \( c \) and for \( n > 2 \), it satisfies the recursive equation

\[ T(n) \leq 3T(\lfloor n/2 \rfloor + 1) + c'n \quad (5) \]

for some constant \( c' \) (using the fact that addition can be done in \( O(n) \) operations).

The recursive equation (5) solves to \( O(n \log_3^3) \). The intuition behind this is presented in Fig. 2, and this is also a consequence of the so-called “Master Theorem” on recurrence relations that is covered in many discrete math texts. As mentioned above, we leave completing the proof to the reader as Exercise 0.4.

\[ \text{Figure 4: Karatsuba’s algorithm reduces an} \ n \ \text{-bit multiplication to three} \ n/2 \ \text{-bit multiplications, which in turn are reduced to nine} \ n/4 \ \text{-bit multiplications and so on. We can represent the computational cost of all these multiplications in a 3-ary tree of depth} \ \log_2 n, \ \text{where at the root the extra cost is} \ cn \ \text{operations, at the first level the extra cost is} \ c(n/2) \ \text{operations, and at each of the} \ 3^i \ \text{nodes of level} \ i, \ \text{the extra cost is} \ c(n/2^i). \ \text{The total cost is} \ cn \sum_{i=0}^{\log_2 n} (3/2)^i \leq 10cn \log_3^3 \ \text{by the formula for summing a geometric series.} \]
Karatsuba’s algorithm is by no means the end of the line for multiplication algorithms. In the 1960’s, Toom and Cook extended Karatsuba’s ideas to get an $O(n \log_k (2k-1))$ time multiplication algorithm for every constant $k$. In 1971, Schönhage and Strassen got even better algorithms using the Fast Fourier Transform; their idea was to somehow treat integers as “signals” and do the multiplication more efficiently by moving to the Fourier domain. (The Fourier transform is a central tool in mathematics and engineering, used in a great many applications; if you have not seen it yet, you are likely encounter it at some point in your studies.) In the years that followed researchers kept improving the algorithm, and only very recently Harvey and Van Der Hoeven managed to obtain an $O(n \log n)$ time algorithm for multiplication (though it only starts beating the Schönhage-Strassen algorithm for truly astronomical numbers). Yet, despite all this progress, we still don’t know whether or not there is an $O(n)$ time algorithm for multiplying two $n$ digit numbers!

**Remark 0.7 — Matrix Multiplication (advanced note).**

(This book contains many “advanced” or “optional” notes and sections. These may assume background that not every student has, and can be safely skipped over as none of the future parts depends on them.)

Ideas similar to Karatsuba’s can be used to speed up matrix multiplications as well. Matrices are a powerful way to represent linear equations and operations, widely used in a great many applications of scientific computing, graphics, machine learning, and many many more.

One of the basic operations one can do with two matrices is to multiply them. For example, if

$$x = \begin{pmatrix} x_{0,0} & x_{0,1} \\ x_{1,0} & x_{1,1} \end{pmatrix} \quad \text{and} \quad y = \begin{pmatrix} y_{0,0} & y_{0,1} \\ y_{1,0} & y_{1,1} \end{pmatrix}.$$ 

then the product of $x$ and $y$ is the matrix

$$\begin{pmatrix} x_{0,0}y_{0,0} + x_{0,1}y_{1,0} & x_{0,0}y_{0,1} + x_{0,1}y_{1,1} \\ x_{1,0}y_{0,0} + x_{1,1}y_{1,0} & x_{1,0}y_{0,1} + x_{1,1}y_{1,1} \end{pmatrix}.$$ 

You can see that we can compute this matrix by eight products of numbers.

Now suppose that $n$ is even and $x$ and $y$ are a pair of $n \times n$ matrices which we can think of as each composed of four $(n/2) \times (n/2)$ blocks $x_{0,0}, x_{0,1}, x_{1,0}, x_{1,1}$ and $y_{0,0}, y_{0,1}, y_{1,0}, y_{1,1}$. Then the formula for the matrix product of $x$ and $y$ can be expressed in the same way as above, just replacing products $x_{a,b}y_{c,d}$ with matrix products, and addition with matrix addition. This means that we can use the formula above to give an algorithm that doubles the dimension of the matrices at the expense of increasing the number of operation
by a factor of 8, which for \( n = 2^\ell \) results in \( 8^\ell = n^3 \) operations.

In 1969 Volker Strassen noted that we can compute the product of a pair of two-by-two matrices using only seven products of numbers by observing that each entry of the matrix \( xy \) can be computed by adding and subtracting the following seven terms:

\[
\begin{align*}
 t_1 &= (x_{0,0} + x_{1,1})(y_{0,0} + y_{1,1}),
 t_2 &= (x_{0,0} + x_{1,1})y_{0,0},
 t_3 &= x_{0,0}(y_{0,1} - y_{1,1}),
 t_4 &= x_{1,1}(y_{0,1} - y_{0,0}),
 t_5 &= (x_{0,0} + x_{0,1})y_{1,1},
 t_6 &= (x_{1,0} - x_{0,0})(y_{0,0} + y_{0,1}),
 t_7 &= (x_{0,1} - x_{1,1})(y_{1,0} + y_{1,1}).
\end{align*}
\]

Indeed, one can verify that \( xy = (t_1 + t_4 - t_5 + t_7 - t_3 + t_5)
\]

Using this observation, we can obtain an algorithm such that doubling the dimension of the matrices results in increasing the number of operations by a factor of 7, which means that for \( n = 2^\ell \) the cost is \( 7^\ell = n^{\log_2 7} \sim n^{2.807} \). A long sequence of work has since improved this algorithm, and the current record has running time about \( O(n^{2.373}) \). However, unlike the case of integer multiplication, at the moment we don't know of any algorithm for matrix multiplication that runs in time linear or even close to linear in the size of the input matrices (e.g., an \( O(n^{2 \text{polylog}(n)}) \) time algorithm). People have tried to use group representations, which can be thought of as generalizations of the Fourier transform, to obtain faster algorithms, but this effort has not yet succeeded.

### 0.3 Algorithms Beyond Arithmetic

The quest for better algorithms is by no means restricted to arithmetic tasks such as adding, multiplying or solving equations. Many graph algorithms, including algorithms for finding paths, matchings, spanning trees, cuts, and flows, have been discovered in the last several decades, and this is still an intensive area of research. (For example, the last few years saw many advances in algorithms for the maximum flow problem, borne out of unexpected connections with electrical circuits and linear equation solvers.) These algorithms are being used not just for the “natural” applications of routing network traffic or GPS-based navigation, but also for applications as varied as drug discovery through searching for structures in gene-interaction graphs to computing risks from correlations in financial investments.

Google was founded based on the PageRank algorithm, which is an efficient algorithm to approximate the “principal eigenvector” of (a dampened version of) the adjacency matrix of the web graph. The Akamai company was founded based on a new data structure, known as consistent hashing, for a hash table where buckets are stored at dif-


ficient servers. The backpropagation algorithm, which computes partial
derivatives of a neural network in $O(n)$ instead of $O(n^2)$ time, under-
lies many of the recent phenomenal successes of learning deep neural
networks. Algorithms for solving linear equations under sparsity
constraints, a concept known as compressed sensing, have been used
to drastically reduce the amount and quality of data needed to ana-
lyze MRI images. This made a critical difference for MRI imaging of
cancer tumors in children, where previously doctors needed to use
anesthesia to suspend breath during the MRI exam, sometimes with
dire consequences.

Even for classical questions, studied through the ages, new dis-
coveries are still being made. For example, for the question of de-
termining whether a given integer is prime or composite, which has
been studied since the days of Pythagoras, efficient probabilistic algo-
rithms were only discovered in the 1970s, while the first deterministic
polynomial-time algorithm was only found in 2002. For the related
problem of actually finding the factors of a composite number, new
algorithms were found in the 1980s, and (as we’ll see later in this
course) discoveries in the 1990s raised the tantalizing prospect of
obtaining faster algorithms through the use of quantum mechanical
effects.

Despite all this progress, there are still many more questions than
answers in the world of algorithms. For almost all natural prob-
lems, we do not know whether the current algorithm is the “best”,
or whether a significantly better one is still waiting to be discovered.
As alluded in Cobham’s opening quote for this chapter, even for the
basic problem of multiplying numbers we have not yet answered the
question of whether there is a multiplication algorithm that is as ef-
ficient as our algorithms for addition. But at least we now know the
right way to ask it.

0.4 ON THE IMPORTANCE OF NEGATIVE RESULTS.

Finding better algorithms for problems such as multiplication, solv-
ing equations, graph problems, or fitting neural networks to data, is
undoubtedly a worthwhile endeavor. But why is it important to prove
that such algorithms don’t exist? One motivation is pure intellectual
curiosity. Another reason to study impossibility results is that they
correspond to the fundamental limits of our world. In other words,
impossibility results are laws of nature.

Here are some examples of impossibility results outside computer
science (see Section 0.7 for more about these). In physics, the impos-
sibility of building a perpetual motion machine corresponds to the law
of conservation of energy. The impossibility of building a heat engine
beating Carnot’s bound corresponds to the second law of thermo-
dynamics, while the impossibility of faster-than-light information transmission is a cornerstone of special relativity. In mathematics, while we all learned the formula for solving quadratic equations in high school, the impossibility of generalizing this formula to equations of degree five or more gave birth to group theory. The impossibility of proving Euclid’s fifth axiom from the first four gave rise to non-Euclidean geometries, which ended up crucial for the theory of general relativity.

In an analogous way, impossibility results for computation correspond to “computational laws of nature” that tell us about the fundamental limits of any information processing apparatus, whether based on silicon, neurons, or quantum particles. Moreover, computer scientists found creative approaches to apply computational limitations to achieve certain useful tasks. For example, much of modern Internet traffic is encrypted using the RSA encryption scheme, which relies on its security on the (conjectured) impossibility of efficiently factoring large integers. More recently, the Bitcoin system uses a digital analog of the “gold standard” where, instead of using a precious metal, new currency is obtained by “mining” solutions for computationally difficult problems.

Lecture Recap

- The history of algorithms goes back thousands of years; they have been essential much of human progress and these days form the basis of multi-billion dollar industries, as well as life-saving technologies.
- There is often more than one algorithm to achieve the same computational task. Finding a faster algorithm can often make a much bigger difference than improving computing hardware.
- Better algorithms and data structures don’t just speed up calculations, but can yield new qualitative insights.
- One question we will study is to find out what is the most efficient algorithm for a given problem.
- To show that an algorithm is the most efficient one for a given problem, we need to be able to prove that it is impossible to solve the problem using a smaller amount of computational resources.

0.5 ROADMAP TO THE REST OF THIS BOOK

Often, when we try to solve a computational problem, whether it is solving a system of linear equations, finding the top eigenvector of a matrix, or trying to rank Internet search results, it is enough to use the
“I know it when I see it” standard for describing algorithms. As long as we find some way to solve the problem, we are happy and might not care much on the exact mathematical model for our algorithm. But when we want to answer a question such as “does there exist an algorithm to solve the problem $P$?” we need to be much more precise.

In particular, we will need to (1) define exactly what it means to solve $P$, and (2) define exactly what an algorithm is. Even (1) can sometimes be non-trivial but (2) is particularly challenging; it is not at all clear how (and even whether) we can encompass all potential ways to design algorithms. We will consider several simple models of computation, and argue that, despite their simplicity, they do capture all “reasonable” approaches to achieve computing, including all those that are currently used in modern computing devices.

Once we have these formal models of computation, we can try to obtain impossibility results for computational tasks, showing that some problems cannot be solved (or perhaps cannot be solved within the resources of our universe). Archimedes once said that given a fulcrum and a long enough lever, he could move the world. We will see how reductions allow us to leverage one hardness result into a slew of a great many others, illuminating the boundaries between the computable and uncomputable (or tractable and intractable) problems.

Later in this book we will go back to examining our models of computation, and see how resources such as randomness or quantum entanglement could potentially change the power of our model. In the context of probabilistic algorithms, we will see a glimpse of how randomness has become an indispensable tool for understanding computation, information, and communication. We will also see how computational difficulty can be an asset rather than a hindrance, and be used for the “derandomization” of probabilistic algorithms. The same ideas also show up in cryptography, which has undergone not just a technological but also an intellectual revolution in the last few decades, much of it building on the foundations that we explore in this course.

Theoretical Computer Science is a vast topic, branching out and touching upon many scientific and engineering disciplines. This book provides a very partial (and biased) sample of this area. More than anything, I hope I will manage to “infect” you with at least some of my love for this field, which is inspired and enriched by the connection to practice, but is also deep and beautiful regardless of applications.

0.5.1 Dependencies between chapters
This book is divided into the following parts, see Fig. 5.
• **Preliminaries**: Introduction, mathematical background, and representing objects as strings.

• **Part I: Finite computation (Boolean circuits)**: Equivalence of circuits and straight-line programs. Universal gate sets. Existence of a circuit for every function, representing circuits as strings, universal circuit, lower bound on circuit size using the counting argument.

• **Part II: Uniform computation (Turing machines)**: Equivalence of Turing machines and programs with loops. Equivalence of models (including RAM machines, $\lambda$ calculus, and cellular automata), configurations of Turing machines, existence of a universal Turing machine, uncomputable functions (including the Halting problem and Rice’s Theorem), Gödel’s incompleteness theorem, restricted computational models models (regular and context free languages).

• **Part III: Efficient computation**: Definition of running time, time hierarchy theorem, $P$ and $NP$, $P_{/\text{poly}}$, $NP$ completeness and the Cook-Levin Theorem, space bounded computation.

• **Part IV: Randomized computation**: Probability, randomized algorithms, $BPP$, amplification, $BPP \subseteq P_{/\text{poly}}$, pseudorandom generators and derandomization.

• **Part V: Advanced topics**: Cryptography, proofs and algorithms (interactive and zero knowledge proofs, Curry-Howard correspondence), quantum computing.

**Figure 5**: The dependency structure of the different parts. Part I introduces the model of Boolean circuits to study finite functions with an emphasis on quantitative questions (how many gates to compute a function). Part II introduces the model of Turing machines to study functions that have unbounded input lengths with an emphasis on qualitative questions (is this function computable or not). Much of Part II does not depend on Part I, as Turing machines can be used as the first computational model. Part III depends on both parts as it introduces a quantitative study of functions with unbounded input length. The more advanced parts IV (randomized computation) and V (advanced topics) rely on the material of Parts I, II and III.

The book largely proceeds in linear order, with each chapter building on the previous ones, with the following exceptions:
• The topics of λ calculus (Section 7.5 and Section 7.5), Gödel’s incompleteness theorem (Chapter 10), Automata/regular expressions and context-free grammars (Chapter 9), and space-bounded computation (Chapter 16), are not used in the following chapters. Hence you can choose whether to cover or skip any subset of them.

• Part II (Uniform Computation / Turing Machines) does not have a strong dependency on Part I (Finite computation / Boolean circuits) and it should be possible to teach them in the reverse order with minor modification. Boolean circuits are used Part III (efficient computation) for results such as $P \subseteq P_{/\text{poly}}$ and the Cook-Levin Theorem, as well as in Part IV (for $\text{BPP} \subseteq P_{/\text{poly}}$ and derandomization) and Part V (specifically in cryptography and quantum computing).

• All chapters in Part V (Advanced topics) are independent of one another and can be covered in any order.

A course based on this book can use all of Parts I, II, and III (possibly skipping over some or all of the λ calculus, Chapter 10, Chapter 9 or Chapter 16), and then either cover all or some of Part IV (randomized computation), and add a “sprinkling” of advanced topics from Part V based on student or instructor interest.

0.6 EXERCISES

Exercise 0.1 Rank the significance of the following inventions in speeding up multiplication of large (that is 100-digit or more) numbers. That is, use “back of the envelope” estimates to order them in terms of the speedup factor they offered over the previous state of affairs.

a. Discovery of the grade-school digit by digit algorithm (improving upon repeated addition)

b. Discovery of Karatsuba’s algorithm (improving upon the digit by digit algorithm)

c. Invention of modern electronic computers (improving upon calculations with pen and paper).

Exercise 0.2 The 1977 Apple II personal computer had a processor speed of 1.023 Mhz or about $10^6$ operations per seconds. At the time of this writing the world’s fastest supercomputer performs 93 “petaflops” ($10^{15}$ floating point operations per second) or about $10^{18}$ basic steps per second. For each one of the following running times (as a function of the input length $n$), compute for both computers how
large an input they could handle in a week of computation, if they run an algorithm that has this running time:

a. \( n \) operations.
b. \( n^2 \) operations.
c. \( n \log n \) operations.
d. \( 2^n \) operations.
e. \( n! \) operations.

**Exercise 0.3 — Usefulness of algorithmic non-existence.** In this chapter we mentioned several companies that were founded based on the discovery of new algorithms. Can you give an example for a company that was founded based on the non existence of an algorithm? See footnote for hint.4

**Exercise 0.4 — Analysis of Karatsuba’s Algorithm.**

a. Suppose that \( T_1, T_2, T_3, \ldots \) is a sequence of numbers such that \( T_2 \leq 10 \) and for every \( n, T_n \leq 3T_{\lfloor n/2 \rfloor} + 1 + C \) \( n \) for some \( C \geq 1 \). Prove that \( T_n \leq 20 C n \log_3^3 \) for every \( n > 2 \).5

b. Prove that the number of single-digit operations that Karatsuba’s algorithm takes to multiply two \( n \) digit numbers is at most \( 1000 n \log_3^3 \).

**Exercise 0.5** Implement in the programming language of your choice functions `Gradeschool_multiply(x, y)` and `Karat$\theta$uba_multiply(x, y)` that take two arrays of digits \( x \) and \( y \) and return an array representing the product of \( x \) and \( y \) (where \( x \) is identified with the number \( x[0] + 10x[1] + 100x[2] + \ldots \) etc..) using the grade-school algorithm and the Karatsuba algorithm respectively. At what number of digits does the Karatsuba algorithm beat the grade-school one?

**Exercise 0.6 — Matrix Multiplication (optional, advanced).** In this exercise, we show that if for some \( \omega > 2 \), we can write the product of two \( k \times k \) real-valued matrices \( A, B \) using at most \( k^\omega \) multiplications, then we can multiply two \( n \times n \) matrices in roughly \( n^\omega \) time for every large enough \( n \).

To make this precise, we need to make some notation that is unfortunately somewhat cumbersome. Assume that there is some \( k \in \mathbb{N} \)

---

4 As we will see in Chapter 20, almost any company relying on cryptography needs to assume the non existence of certain algorithms. In particular, RSA Security was founded based on the security of the RSA cryptosystem, which presumes the non existence of an efficient algorithm to compute the prime factorization of large integers.

5 Hint: Use a proof by induction - suppose that this is true for all \( n \)’s from 1 to \( m \) and prove that this is true also for \( m + 1 \).
and \( m \leq k^\omega \) such that for every \( k \times k \) matrices \( A, B, C \) such that \( C = AB \), we can write for every \( i, j \in [k] \):

\[
C_{i,j} = \sum_{\ell=0}^{m} \alpha_{i,j}^{\ell} f_{\ell}(A) g_{\ell}(B)
\]

for some linear functions \( f_0, \ldots, f_{m-1}, g_0, \ldots, g_{m-1} : \mathbb{R}^{n^2} \to \mathbb{R} \) and coefficients \( \{ \alpha_{i,j}^{\ell} \}_{i,j \in [k], \ell \in [m]} \). Prove that under this assumption for every \( \epsilon > 0 \), if \( n \) is sufficiently large, then there is an algorithm that computes the product of two \( n \times n \) matrices using at most \( O(n^{\omega + \epsilon}) \) arithmetic operations. See footnote for hint.\(^6\)

0.7 BIBLIOGRAPHICAL NOTES

For a brief overview of what we’ll see in this book, you could do far worse than read Bernard Chazelle’s wonderful essay on the Algorithm as an Idiom of modern science. The book of Moore and Mertens [MM11] gives a wonderful and comprehensive overview of the theory of computation, including much of the content discussed in this chapter and the rest of this book. Aaronson’s book [Aar13] is another great read that touches upon many of the same themes.

For more on the algorithms the Babylonians used, see Knuth’s paper and Neugebauer’s classic book.

Many of the algorithms we mention in this chapter are covered in algorithms textbooks such as those by Cormen, Leiserson, Rivert, and Stein [Cor+09], Kleinberg and Tardos [KT06], and Dasgupta, Papadimitriou and Vazirani [DPV08], as well as Jeff Erickson’s textbook. Erickson’s book is freely available online and contains a great exposition of recursive algorithms in general and Karatsuba’s algorithm in particular.

The story of Karatsuba’s discovery of his multiplication algorithm is recounted by him in [Kar95]. As mentioned above, further improvements were made by Toom and Cook [Too63; Coo66], Schönhage and Strassen [SS71], Fürer [Für07], and recently by Harvey and Van Der Hoeven [HV19], see this article for a nice overview. The last papers crucially rely on the Fast Fourier transform algorithm. The fascinating story of the (re)discovery of this algorithm by John Tukey in the context of the cold war is recounted in [Coo87]. (We say re-discovery because it later turned out that the algorithm dates back to Gauss [HJB85].) The Fast Fourier Transform is covered in some of the books mentioned below, and there are also online available lectures such as Jeff Erickson’s. See also this popular article by David Austin. Fast matrix multiplication was discovered by Strassen [Str69], and since then this has been an active area of research. [Bla13] is a recommended self-contained survey of this area.

\(^6\) Start by showing this for the case that \( n = k^t \) for some natural number \( t \), in which case you can do so recursively by breaking the matrices into \( k \times k \) blocks.
The Backpropagation algorithm for fast differentiation of neural networks was invented by Werbos [Wer74]. The Pagerank algorithm was invented by Larry Page and Sergey Brin [Pag+99]. It is closely related to the HITS algorithm of Kleinberg [Kle99]. The Akamai company was founded based on the consistent hashing data structure described in [Kar+97]. Compressed sensing has a long history but two foundational papers are [CRT06; Don06]. [Lus+08] gives a survey of applications of compressed sensing to MRI; see also this popular article by Ellenberg [Ell10]. The deterministic polynomial-time algorithm for testing primality was given by Agrawal, Kayal, and Saxena [AKS04].

We alluded briefly to classical impossibility results in mathematics, including the impossibility of proving Euclid’s fifth postulate from the other four, impossibility of trisecting an angle with a straightedge and compass and the impossibility of solving a quintic equation via radicals. A geometric proof of the impossibility of angle trisection (one of the three geometric problems of antiquity, going back to the ancient Greeks) is given in this blog post of Tao. The book of Mario Livio [Liv05] covers some of the background and ideas behind these impossibility results. Some exciting recent research is focused on trying to use computational complexity to shed light on fundamental questions in physics such understanding black holes and reconciling general relativity with quantum mechanics.
1

Mathematical Background

“I found that every number, which may be expressed from one to ten, surpasses the preceding by one unit: afterwards the ten is doubled or tripled ... until a hundred; then the hundred is doubled and tripled in the same manner as the units and the tens ... and so forth to the utmost limit of numeration.”, Muhammad ibn Mūsā al-Khwārizmī, 820, translation by Fredric Rosen, 1831.

In this chapter we review some of the mathematical concepts that we use in this book. These concepts are typically covered in courses or textbooks on “mathematics for computer science” or “discrete mathematics”; see the “Bibliographical Notes” section (Section 1.9) for several excellent resources on these topics that are freely-available online.

A mathematician’s apology. Some students might wonder why this book contains so much math. Mathematics is a language for modeling concepts in a precise and unambiguous way. In this book we use math to model the concept of computation. For example, we will consider questions such as “is there an efficient algorithm to find the prime factors of a given integer?”. (We will see that this question is particularly interesting, touching on areas as far apart as Internet security and quantum mechanics!) To even phrase such a question, we need to give a precise definition of the notion of an algorithm, and of what it means for an algorithm to be efficient. Also, since there is no empirical experiment that will prove the nonexistence of an algorithm, the only way to establish such a result is using a mathematical proof.

1.1 THIS CHAPTER: A READER’S MANUAL

Depending on your background, you can approach this chapter in two different ways:
• If you already have taken a “discrete mathematics”, “mathematics for computer science” or similar courses, you can take a quick look at Section 1.2 to see the main tools we will use, Section 1.7 for our notation and conventions, and then skip ahead to the rest of this book. Alternatively, you can sit back, relax, and read this chapter just to get familiar with our notation, as well as to enjoy (or not) my philosophical musings and attempts at humor. You might also want to start brushing up on discrete probability, which we’ll use later in this book.

• If your background is less extensive, see Section 1.9 for some resources on these topics. This chapter briefly covers the concepts that we need, but you may find it helpful to see a more in-depth treatment. As usual with math, the best way to get comfort with this material is to work out exercises on your own.

1.2 A QUICK OVERVIEW OF MATHEMATICAL PREREQUISITES

The main mathematical concepts we use in this book are:

• **Proofs:** First and foremost, this book involves a heavy dose of formal mathematical reasoning, which includes mathematical definitions, statements, and proofs.

• **Sets:** The basic set relations of membership (∈) and containment (⊆), and set operations, principally union (∪), intersection (∩), set difference (\) and Cartesian product (×).

• **Tuples and strings:** The set Σ^k of length-k strings/lists over elements in Σ, where Σ is some finite set which is called the alphabet (quite often Σ = {0, 1}). We use Σ^* for the set of all strings of finite length.

• **Some special sets:** The set ℤ of natural numbers. Following typical computer science convention, our indices start from zero and so we write ℤ = {0, 1, 2, ...}. We use [n] for the set {0, 1, 2, ..., n − 1}. We use {0, 1}^* for the set of all binary strings and {0, 1}^n for the set of strings of length n for some natural number n ∈ ℤ. If x is a string of length n, then we refer to its elements by x_0, ..., x_{n−1}.

• **Functions:** The domain and codomain of a function, properties such as being one-to-one (also known as injective) or onto (also known as surjective) functions, as well as partial functions (that, unlike standard or “total” functions, are not necessarily defined on all elements of their domain).

• **Logical operations:** The operations AND (∧), OR (∨), and NOT (¬) and the quantifiers “there exists” (∃) and “for all” (∀).
• **Basic combinatorics**: Notions such as \( \binom{n}{k} \) (the number of \( k \)-sized subsets of a set of size \( n \)).

• **Graphs**: Undirected and directed graphs, connectivity, paths, and cycles.

• **Big-O notation**: \( O, o, \Omega, \omega, \Theta \) notation for analyzing asymptotic growth of functions.

• **Discrete probability**: We will use probability theory, and specifically probability over finite samples spaces such as tossing \( n \) coins, including notions such as random variables, expectation, and concentration. We will only use probability theory in the second half of this text, and will review it beforehand. However, probabilistic reasoning is a subtle (and extremely useful!) skill, and it’s always good to start early in acquiring it.

In the rest of this chapter we briefly review the above notions. This is partially to remind the reader and reinforce material that might not be fresh in your mind, and partially to introduce our notation and conventions which might occasionally differ from those you’ve encountered before.

### 1.3 READING MATHEMATICAL TEXTS

Reading mathematical texts take practice to get used to the notation and symbols. Mathematicians use jargon for the same reason that it is used in many other professions such as engineering, law, medicine, and others. We want to make terms precise and introduce shorthand for concepts that are frequently reused. Mathematical texts tend to “pack a lot of punch” per sentence, and so the key is to read them slowly and carefully, parsing each symbol at a time.

With time and practice you will see that reading mathematical texts becomes easier and jargon is no longer an issue. Moreover, reading mathematical texts is one of the most transferable skills you could take from this book. Our world is changing rapidly, not just in the realm of technology, but also in many other human endeavors, whether it is medicine, economics, law or even culture. Whatever your future aspirations, it is likely that you will encounter texts that use new concepts that you have not seen before (see Fig. 1.1 and Fig. 1.2 for two recent examples from current “hot areas”). Being able to internalize and then apply new definitions can be hugely important. It is a skill that’s much easier to acquire in the relatively safe and stable context of a mathematical course, where one at least has the guarantee that the concepts are fully specified, and you have access to your teaching staff for questions.
The basic components of a mathematical text are definitions, assertions and proofs.

1.3.1 Definitions
Mathematicians often define new concepts in terms of old concepts.

Here is a mathematical definition which you may have encountered in the past (and will see again shortly):

Definition 1.1 — One to one function. Let $S$, $T$ be sets. We say that a function $f : S \to T$ is one to one (also known as injective) if for every two elements $x, x' \in S$, if $x \neq x'$ then $f(x) \neq f(x')$.

Definition 1.1 captures a simple concept, but even so it uses quite a bit of notation. When reading such a definition, it is often useful to annotate it with a pen as you’re going through it, as in Fig. 1.3. For example, when you see an identifier such as $f$, $S$ or $x$, make sure that you realize what sort of object is it: is it a set, a function, an element, a number, a gremlin? You might also find it useful to explain the definition in words to a friend (or to yourself).

1.3.2 Assertions: Theorems, lemmas, claims
Theorems, lemmas, claims and the like are true statements about the concepts that we defined. Deciding whether to call a particular statement a “Theorem”, a “Lemma” or a “Claim” is a judgement call, and does not make a mathematical difference. All three correspond to true statements which can be proven. The difference is that a Theorem refers to a significant result, that we would want to remember and highlight. A Lemma often refers to a technical result, that is not necessarily important in its own right, but can be often very useful in proving other theorems. A Claim is a “throw away” statement, that we need to use in order to prove some other bigger results, but do not care so much about for its own sake.

1.3.3 Proofs
Mathematical proofs are the arguments we use to demonstrate that our theorems, lemmas, and claims area indeed true. We discuss proofs in Section 1.5 below, but the main point is that the mathematical standard of proof is very high. Unlike in some other realms, in mathematics a proof is an “airtight” argument that demonstrates that the statement is true beyond a shadow of a doubt. Some examples in this section for mathematical proofs are given in Solved Exercise 1.1 and Section 1.6. As mentioned in the preface, as a general rule, it is more important you understand the definitions than the theorems, and it is more important you understand a theorem statement than its proof.
1.4 BASIC DISCRETE MATH OBJECTS

In this section we quickly review some of the mathematical objects (the “basic data structures” of mathematics, if you will) we use in this book.

1.4.1 Sets

A set is an unordered collection of objects. For example, when we write \( S = \{2, 4, 7\} \), we mean that \( S \) denotes the set that contains the numbers 2, 4, and 7. (We use the notation “\( 2 \in S \)” to denote that 2 is an element of \( S \).) Note that the set \( \{2, 4, 7\} \) and \( \{7, 4, 2\} \) are identical, since they contain the same elements. Also, a set either contains an element or does not contain it – there is no notion of containing it “twice” – and so we could even write the same set \( S \) as \( \{2, 2, 4, 7\} \) (though that would be a little weird). The cardinality of a finite set \( S \), denoted by \(|S|\), is the number of elements it contains. (Cardinality can be defined for infinite sets as well; see the sources in Section 1.9.) So, in the example above, \(|S| = 3\). A set \( S \) is a subset of a set \( T \), denoted by \( S \subseteq T \), if every element of \( S \) is also an element of \( T \). (We can also describe this by saying that \( T \) is a superset of \( S \).) For example, \( \{2, 7\} \subseteq \{2, 4, 7\} \). The set that contains no elements is known as the empty set and it is denoted by \( \emptyset \).

We can define sets by either listing all their elements or by writing down a rule that they satisfy such as

\[
\text{EVEN} = \{x \mid x = 2y \text{ for some non-negative integer } y\} \quad (1.1)
\]

Of course there is more than one way to write the same set, and often we will use intuitive notation listing a few examples that illustrate the rule. For example, we can also define EVEN as

\[
\text{EVEN} = \{0, 2, 4, ...\} \quad (1.2)
\]

Note that a set can be either finite (such as the set \( \{2, 4, 7\} \)) or infinite (such as the set \( \text{EVEN} \)). Also, the elements of a set don’t have to be numbers. We can talk about the sets such as the set \( \{a, e, i, o, u\} \) of all the vowels in the English language, or the set \( \{\text{New York, Los Angeles, Chicago, Houston, Philadelphia, Phoenix, San Antonio, San Diego, Dallas}\} \) of all cities in the U.S. with population more than one million per the 2010 census. A set can even have other sets as elements, such as the set \( \{\emptyset, \{1, 2\}, \{2, 3\}, \{1, 3\}\} \) of all even-sized subsets of \( \{1, 2, 3\} \).

**Operations on sets:** The union of two sets \( S, T \), denoted by \( S \cup T \), is the set that contains all elements that are either in \( S \) or in \( T \). The intersection of \( S \) and \( T \), denoted by \( S \cap T \), is the set of elements that are both in \( S \) and in \( T \). The set difference of \( S \) and \( T \), denoted by \( S \setminus T \) (and
in some texts also by $S - T$), is the set of elements that are in $S$ but not in $T$.

**Tuples, lists, strings, sequences:** A tuple is an ordered collection of items. For example $(1, 5, 2, 1)$ is a tuple with four elements (also known as a 4-tuple or quadruple). Since order matters, this is not the same tuple as the 4-tuple $(1, 1, 5, 2)$ or the 3-tuple $(1, 5, 2)$. A 2-tuple is also known as a pair. We use the terms tuples and lists interchangeably.

A tuple where every element comes from some finite set $\Sigma$ (such as $\{0, 1\}$) is also known as a string. Analogously to sets, we denote the length of a tuple $T$ by $|T|$. Just like sets, we can also think of infinite analogues of tuples, such as the ordered collection $(1, 4, 9, \ldots)$ of all perfect squares. Infinite ordered collections are known as sequences; we might sometimes use the term “infinite sequence” to emphasize this, and use “finite sequence” as a synonym for a tuple. (We can identify a sequence $(a_0, a_1, a_2, \ldots)$ of elements in some set $S$ with a function $A : \mathbb{N} \to S$ (where $a_n = A(n)$ for every $n \in \mathbb{N}$). Similarly, we can identify a $k$-tuple $(a_0, \ldots, a_{k-1})$ of elements in $S$ with a function $A : [k] \to S$.)

**Cartesian product:** If $S$ and $T$ are sets, then their Cartesian product, denoted by $S \times T$, is the set of all ordered pairs $(s, t)$ where $s \in S$ and $t \in T$. For example, if $S = \{1, 2, 3\}$ and $T = \{10, 12\}$, then $S \times T$ contains the 6 elements $(1, 10), (2, 10), (3, 10), (1, 12), (2, 12), (3, 12)$. Similarly if $S, T, U$ are sets then $S \times T \times U$ is the set of all ordered triples $(s, t, u)$ where $s \in S, t \in T$, and $u \in U$. More generally, for every positive integer $n$ and sets $S_0, \ldots, S_{n-1}$, we denote by $S_0 \times S_1 \times \cdots \times S_{n-1}$ the set of ordered $n$-tuples $(s_0, \ldots, s_{n-1})$ where $s_i \in S_i$ for every $i \in \{0, \ldots, n-1\}$. For every set $S$, we denote the set $S \times S$ by $S^2$, $S \times S \times S$ by $S^3$, $S \times S \times S \times S$ by $S^4$, and so on and so forth.

**1.4.2 Special sets**

There are several sets that we will use in this book time and again. The set

$$\mathbb{N} = \{0, 1, 2, \ldots\}$$

(1.3)

contains all natural numbers, i.e., non-negative integers. For any natural number $n \in \mathbb{N}$, we define the set $[n]$ as $\{0, \ldots, n-1\} = \{k \in \mathbb{N} : k < n\}$. (We start our indexing of both $\mathbb{N}$ and $[n]$ from 0, while many other texts index those sets from 1. Starting from zero or one is simply a convention that doesn’t make much difference, as long as one is consistent about it.)

We will also occasionally use the set $\mathbb{Z} = \{\ldots, -2, -1, 0, +1, +2, \ldots\}$ of (negative and non-negative) integers, as well as the set $\mathbb{R}$ of real numbers. (This is the set that includes not just the integers, but also

\[^1\text{The letter Z stands for the German word “Zahlen”, which means numbers.}\]
fractional and irrational numbers; e.g., \( \mathbb{R} \) contains numbers such as \(+0.5, -\pi, \text{etc.}\) We denote by \( \mathbb{R}_+ \), the set \( \{ x \in \mathbb{R} : x > 0 \} \) of positive real numbers. This set is sometimes also denoted as \((0, \infty)\).

**Strings:** Another set we will use time and again is

\[
\{0,1\}^n = \{(x_0, \ldots, x_{n-1}) : x_0, \ldots, x_{n-1} \in \{0, 1\}\} 
\]

which is the set of all \( n \)-length binary strings for some natural number \( n \). That is \( \{0,1\}^n \) is the set of all \( n \)-tuples of zeroes and ones. This is consistent with our notation above: \( \{0,1\}^2 \) is the Cartesian product \( \{0,1\} \times \{0,1\} \), \( \{0,1\}^3 \) is the product \( \{0,1\} \times \{0,1\} \times \{0,1\} \) and so on.

We will write the string \((x_0, x_1, \ldots, x_{n-1})\) as simply \(x_0 x_1 \ldots x_{n-1}\). For example,

\[
\{0,1\}^3 = \{000, 001, 010, 011, 100, 101, 110, 111\}.
\]

For every string \( x \in \{0,1\}^n \) and \( i \in [n] \), we write \( x_i \) for the \( i \)th element of \( x \).

We will also often talk about the set of binary strings of all lengths, which is

\[
\{0,1\}^* = \{(x_0, \ldots, x_{n-1}) : n \in \mathbb{N} \cup \{0\}, x_0, \ldots, x_{n-1} \in \{0, 1\}\}.
\]

Another way to write this set is as

\[
\{0,1\}^* = \{0,1\}^0 \cup \{0,1\}^1 \cup \{0,1\}^2 \cup \ldots
\]

or more concisely as

\[
\{0,1\}^* = \cup_{n \in \mathbb{N}} \{0,1\}^n.
\]

The set \( \{0,1\}^* \) includes the “string of length 0” or “the empty string”, which we will denote by “\(\epsilon\)”. (In using this notation we follow the convention of many programming languages. Other texts sometimes use \(\epsilon\) or \(\lambda\) to denote the empty string.)

**Generalizing the star operation:** For every set \( \Sigma \), we define

\[
\Sigma^* = \cup_{n \in \mathbb{N}} \Sigma^n.
\]

For example, if \( \Sigma = \{a, b, c, d, \ldots, z\} \) then \( \Sigma^* \) denotes the set of all finite length strings over the alphabet a-z.

**Concatenation:** The concatenation of two strings \( x \in \Sigma^n \) and \( y \in \Sigma^m \) is the \((n + m)\)-length string \( xy \) obtained by writing \( y \) after \( x \). That is, if \( x \in \{0,1\}^n \) and \( y \in \{0,1\}^m \), then \( xy \) is equal to the string \( z \in \{0,1\}^{n+m} \) such that for \( i \in [n] \), \( z_i = x_i \) and for \( i \in [n, n + m - 1] \), \( z_i = y_{i-n} \).
### 1.4.3 Functions

If $S$ and $T$ are nonempty sets, a function $F$ mapping $S$ to $T$, denoted by $F : S \rightarrow T$, associates with every element $x \in S$ an element $F(x) \in T$. The set $S$ is known as the domain of $F$ and the set $T$ is known as the codomain of $F$. The image of a function $F$ is the set \{ $F(x)$ $|$ $x \in S$ \} which is the subset of $F$’s codomain consisting of all output elements that are mapped from some input. (Some texts use range to denote the image of a function, while other texts use range to denote the codomain of a function. Hence we will avoid using the term “range” altogether.) As in the case of sets, we can write a function either by listing the table of all the values it gives for elements in $S$ or by using a rule. For example if $S = \{0, 1, 2, 3, 4, 5, 6, 7, 8, 9\}$ and $T = \{0, 1\}$, then the table below defines a function $F : S \rightarrow T$. Note that this function is the same as the function defined by the rule $F(x) = (x \text{ mod } 2)$.

<table>
<thead>
<tr>
<th>Input</th>
<th>Output</th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>0</td>
</tr>
<tr>
<td>1</td>
<td>1</td>
</tr>
<tr>
<td>2</td>
<td>0</td>
</tr>
<tr>
<td>3</td>
<td>1</td>
</tr>
<tr>
<td>4</td>
<td>0</td>
</tr>
<tr>
<td>5</td>
<td>1</td>
</tr>
<tr>
<td>6</td>
<td>0</td>
</tr>
<tr>
<td>7</td>
<td>1</td>
</tr>
<tr>
<td>8</td>
<td>0</td>
</tr>
<tr>
<td>9</td>
<td>1</td>
</tr>
</tbody>
</table>

Table 1.1: An example of a function.

If $f : S \rightarrow T$ satisfies that $f(x) \neq f(y)$ for all $x \neq y$ then we say that $f$ is one-to-one (Definition 1.1, also known as an injective function or simply an injection). If $F$ satisfies that for every $y \in T$ there is some $x \in S$ such that $F(x) = y$ then we say that $F$ is onto (also known as a surjective function or simply a surjection). A function that is both one-to-one and onto is known as a bijective function or simply a bijection. A bijection from a set $S$ to itself is also known as a permutation of $S$. If $F : S \rightarrow T$ is a bijection then for every $y \in T$ there is a unique $x \in S$ such that $F(x) = y$. We denote this value $x$ by $F^{-1}(y)$. Note that $F^{-1}$ is itself a bijection from $T$ to $S$ (can you see why?).

Giving a bijection between two sets is often a good way to show they have the same size. In fact, the standard mathematical definition of the notion that “$S$ and $T$ have the same cardinality” is that there

---

2 For two natural numbers $x$ and $a$, $x \equiv a \mod a$ (short-hand for “modulo”) denotes the remainder of $x$ when it is divided by $a$. That is, it is the number $r$ in \{ $0, \ldots, a - 1$ \} such that $x = ak + r$ for some integer $k$. We sometimes also use the notation $x \equiv y \mod a$ to denote the assertion that $x \mod a$ is the same as $y \mod a$. 

exists a bijection $f : S \to T$. Further, the cardinality of a set $S$ is defined to be $n$ if there is a bijection from $S$ to the set $\{0, \ldots, n-1\}$. As we will see later in this book, this is a definition that generalizes to defining the cardinality of infinite sets.

**Partial functions:** We will sometimes be interested in partial functions from $S$ to $T$. A partial function is allowed to be undefined on some subset of $S$. That is, if $F$ is a partial function from $S$ to $T$, then for every $s \in S$, either there is (as in the case of standard functions) an element $F(s)$ in $T$, or $F(s)$ is undefined. For example, the partial function $F(x) = \sqrt{x}$ is only defined on non-negative real numbers. When we want to distinguish between partial functions and standard (i.e., non-partial) functions, we will call the latter total functions. When we say “function” without any qualifier then we mean a total function.

The notion of partial functions is a strict generalization of functions, and so every function is a partial function, but not every partial function is a function. (That is, for every nonempty $S$ and $T$, the set of partial functions from $S$ to $T$ is a proper superset of the set of total functions from $S$ to $T$.) When we want to emphasize that a function $f$ from $A$ to $B$ might not be total, we will write $f : A \rightarrow_p B$. We can think of a partial function $F$ from $S$ to $T$ also as a total function from $S$ to $T \cup \{\bot\}$ where $\bot$ is a special “failure symbol”. So, instead of saying that $F$ is undefined at $x$, we can say that $F(x) = \bot$.

**Basic facts about functions:** Verifying that you can prove the following results is an excellent way to brush up on functions:

- If $F : S \to T$ and $G : T \to U$ are one-to-one functions, then their composition $H : S \to U$ defined as $H(s) = G(F(s))$ is also one to one.

- If $F : S \to T$ is one to one, then there exists an onto function $G : T \to S$ such that $G(F(s)) = s$ for every $s \in S$.

- If $G : T \to S$ is onto then there exists a one-to-one function $F : S \to T$ such that $G(F(s)) = s$ for every $s \in S$.

- If $S$ and $T$ are finite sets then the following conditions are equivalent to one another: (a) $|S| \leq |T|$, (b) there is a one-to-one function $F : S \to T$, and (c) there is an onto function $G : T \to S$. (This is actually true even for infinite $S$ and $T$: in that case (b) (or equivalently (c)) is the commonly accepted definition for $|S| \leq |T|$.)

You can find the proofs of these results in many discrete math texts, including for example, Section 4.5 in the Lehman-Leighton-Meyer notes. However, I

![Figure 1.4: We can represent finite functions as a directed graph where we put an edge from $x$ to $f(x)$. The onto condition corresponds to requiring that every vertex in the codomain of the function has in-degree at least one. The one-to-one condition corresponds to requiring that every vertex in the codomain of the function has in-degree at most one. In the examples above $F$ is an onto function, $G$ is one to one, and $H$ is neither onto nor one to one.](image)
It is possible, and sometimes useful, to think of an undirected graph as the special case of a directed graph that has the special property that for every pair \( u, v \) either both the edges \( (u, v) \) and \( (v, u) \) are present or neither of them is. However, in many settings there is a significant difference between undirected and directed graphs, and so it’s typically best to think of them as separate categories.

Figure 1.5: An example of an undirected and a directed graph. The undirected graph has vertex set \{1, 2, 3, 4\} and edge set \{\{1, 2\}, \{2, 3\}, \{2, 4\}\}. The directed graph has vertex set \{a, b, c\} and the edge set \{(a, b), (b, c), (c, a), (a, c)\}.

1.4.4 Graphs

Graphs are ubiquitous in Computer Science, and many other fields as well. They are used to model a variety of data types including social networks, scheduling constraints, road networks, deep neural nets, gene interactions, correlations between observations, and a great many more. Formal definitions of several kinds of graphs are given next, but if you have not seen graphs before in a course, I urge you to read up on them in one of the sources mentioned in Section 1.9.

Graphs come in two basic flavors: undirected and directed.\(^3\)

**Definition 1.3 — Undirected graphs.** An undirected graph \( G = (V, E) \) consists of a set \( V \) of vertices and a set \( E \) of edges. Every edge is a size two subset of \( V \). We say that two vertices \( u, v \in V \) are neighbors, if the edge \( \{u, v\} \) is in \( E \).

Given this definition, we can define several other properties of graphs and their vertices. We define the degree of \( u \) to be the number of neighbors \( u \) has. A path in the graph is a tuple \((u_0, \ldots, u_k) \in V^{k+1}\), for some \( k > 0 \) such that \( u_{i+1} \) is a neighbor of \( u_i \) for every \( i \in [k] \). A simple path is a path \((u_0, \ldots, u_{k-1})\) where all the \( u_i \)’s are distinct. A cycle is a path \((u_0, \ldots, u_k)\) where \( u_0 = u_k \). We say that two vertices \( u, v \in V \) are connected if either \( u = v \) or there is a path from \((u_0, \ldots, u_k)\) where \( u_0 = u \) and \( u_k = v \). We say that the graph \( G \) is connected if every pair of vertices in it is connected.

\(^3\) It is possible, and sometimes useful, to think of an undirected graph as the special case of a directed graph that has the special property that for every pair \( u, v \) either both the edges \( (u, v) \) and \( (v, u) \) are present or neither of them is. However, in many settings there is a significant difference between undirected and directed graphs, and so it’s typically best to think of them as separate categories.
Here are some basic facts about undirected graphs. We give some informal arguments below, but leave the full proofs as exercises (the proofs can be found in many of the resources listed in Section 1.9).

**Lemma 1.4** In any undirected graph $G = (V, E)$, the sum of the degrees of all vertices is equal to twice the number of edges.

**Lemma 1.4** can be shown by seeing that every edge $\{u, v\}$ contributes twice to the sum of the degrees (once for $u$ and the second time for $v$).

**Lemma 1.5** The connectivity relation is *transitive*, in the sense that if $u$ is connected to $v$, and $v$ is connected to $w$, then $u$ is connected to $w$.

**Lemma 1.5** can be shown by simply attaching a path of the form $(u, u_1, u_2, \ldots, u_{k-1}, v)$ to a path of the form $(v, u'_1, \ldots, u'_{k'-1}, w)$ to obtain the path $(u, u_1, \ldots, u_{k-1}, v, u'_1, \ldots, u'_{k'-1}, w)$ that connects $u$ to $w$.

**Lemma 1.6** For every undirected graph $G = (V, E)$ and connected pair $u, v$, the shortest path from $u$ to $v$ is simple. In particular, for every connected pair there exists a simple path that connects them.

**Lemma 1.6** can be shown by “shortcutting” any non simple path from $u$ to $v$ where the same vertex $w$ appears twice to remove it (see Fig. 1.6). It is a good exercise to transforming this intuitive reasoning to a formal proof:

![Figure 1.6](image.png)

*Figure 1.6:* If there is a path from $u$ to $v$ in a graph that passes twice through a vertex $w$ then we can “shortcut” it by removing the loop from $w$ to itself to find a path from $u$ to $v$ that only passes once through $w$.

**Solved Exercise 1.1 — Connected vertices have simple paths.** Prove Lemma 1.6

**Solution:**

The proof follows the idea illustrated in Fig. 1.6. One complication is that there can be more than one vertex that is visited twice by a path, and so “shortcutting” might not necessarily result in a
simple path; we deal with this by looking at a shortest path between $u$ and $v$. Details follow.

Let $G = (V, E)$ be a graph and $u$ and $v$ in $V$ be two connected vertices in $G$. We will prove that there is a simple graph between $u$ and $v$. Let $k$ be the shortest length of a path between $u$ and $v$ and let $P = (u_0, u_1, u_2, \ldots, u_{k-1}, u_k)$ be a $k$-length path from $u$ to $v$ (there can be more than one such path: if so we just choose one of them). (That is $u_0 = u, u_k = v$, and $(u_\ell, u_{\ell+1}) \in E$ for all $\ell \in [k]$.) We claim that $P$ is simple. Indeed, suppose otherwise that there is some vertex $w$ that occurs twice in the path: $w = u_i$ and $w = u_j$ for some $i < j$. Then we can “shortcut” the path $P$ by considering the path $P' = (u_0, u_1, \ldots, u_{i-1}, w, u_{j+1}, \ldots, u_k)$ obtained by taking the first $i$ vertices of $P$ (from $u_0 = 0$ to the first occurrence of $w$) and the last $k - j$ ones (from the vertex $u_{j+1}$ following the second occurrence of $w$ to $u_k = v$). The path $P'$ is a valid path between $u$ and $v$ since every consecutive pair of vertices in it is connected by an edge (in particular, since $w = u_i = u_j$, both $(u_{i-1}, w)$ and $(w, u_{j+1})$ are edges in $E$), but since the length of $P'$ is $k - (j - i) < k$, this contradicts the minimality of $P$.

\begin{remark}
Finding proofs. Solved Exercise 1.1 is a good example of the process of finding a proof. You start by ensuring you understand what the statement means, and then come up with an informal argument why it should be true. You then transform the informal argument into a rigorous proof. This proof need not be very long or overly formal, but should clearly establish why the conclusion of the statement follow from its assumptions.
\end{remark}

The concepts of degrees and connectivity extend naturally to directed graphs, defined as follows.

\begin{definition}
A directed graph $G = (V, E)$ consists of a set $V$ and a set $E \subseteq V \times V$ of ordered pairs of $V$. We sometimes denote the edge $(u, v)$ also as $u \rightarrow v$. If the edge $u \rightarrow v$ is present in the graph then we say that $v$ is an out-neighbor of $u$ and $u$ is an in-neighbor of $v$.

A directed graph might contain both $u \rightarrow v$ and $v \rightarrow u$ in which case $u$ will be both an in-neighbor and an out-neighbor of $v$ and vice versa. The in-degree of $u$ is the number of in-neighbors it has, and the out-degree of $v$ is the number of out-neighbors it has. A path in the
graph is a tuple \((u_0, \ldots, u_k) \in V^{k+1}\), for some \(k > 0\) such that \(u_{i+1}\) is an out-neighbor of \(u_i\) for every \(i \in [k]\). As in the undirected case, a simple path is a path \((u_0, \ldots, u_{k-1})\) where all the \(u_i's\) are distinct and a cycle is a path \((u_0, \ldots, u_k)\) where \(u_0 = u_k\). One type of directed graphs we often care about is directed acyclic graphs or DAGs, which, as their name implies, are directed graphs without any cycles:

**Definition 1.9 — Directed Acyclic Graphs.** We say that \(G = (V, E)\) is a directed acyclic graph (DAG) if it is a directed graph and there does not exist a list of vertices \(u_0, u_1, \ldots, u_k \in V\) such that \(u_0 = u_k\) and for every \(i \in [k]\), the edge \(u_i \rightarrow u_{i+1}\) is in \(E\).

The lemmas we mentioned above have analogs for directed graphs. We again leave the proofs (which are essentially identical to their undirected analogs) as exercises.

**Lemma 1.10** In any directed graph \(G = (V, E)\), the sum of the in-degrees is equal to the sum of the out-degrees, which is equal to the number of edges.

**Lemma 1.11** In any directed graph \(G\), if there is a path from \(u\) to \(v\) and a path from \(v\) to \(w\), then there is a path from \(u\) to \(w\).

**Lemma 1.12** For every directed graph \(G = (V, E)\) and a pair \(u, v\) such that there is a path from \(u\) to \(v\), the shortest path from \(u\) to \(v\) is simple.

**Remark 1.13 — Labeled graphs.** For some applications we will consider labeled graphs, where the vertices or edges have associated labels (which can be numbers, strings, or members of some other set). We can think of such a graph as having an associated (possibly partial) labelling function \(L : V \cup E \rightarrow \mathcal{L}\), where \(\mathcal{L}\) is the set of potential labels. However we will typically not refer explicitly to this labeling function and simply say things such as “vertex \(v\) has the label \(\alpha\).”

### 1.4.5 Logic operators and quantifiers

If \(P\) and \(Q\) are some statements that can be true or false, then \(P\) AND \(Q\) (denoted as \(P \land Q\)) is a statement that is true if and only if both \(P\) and \(Q\) are true, and \(P\) OR \(Q\) (denoted as \(P \lor Q\)) is a statement that is true if and only if either \(P\) or \(Q\) is true. The negation of \(P\), denoted as \(\neg P\) or \(\overline{P}\), is true if and only if \(P\) is false.

Suppose that \(P(x)\) is a statement that depends on some parameter \(x\) (also sometimes known as an unbound variable) in the sense that for every instantiation of \(x\) with a value from some set \(S\), \(P(x)\) is either true or false. For example, \(x > 7\) is a statement that is not a priori
true or false, but becomes true or false whenever we instantiate \( x \) with some real number. We denote by \( \forall x \in S \) \( P(x) \) the statement that is true if and only if \( P(x) \) is true for every \( x \in S \).\footnote{In this book, we place the variable bound by a quantifier in a subscript and so write \( \forall x \in S \) \( P(x) \). Many other texts do not use this subscript notation and so will write the same statement as \( \forall x \in S \), \( P(x) \).} We denote by \( \exists x \in S \) \( P(x) \) the statement that is true if and only if there exists some \( x \in S \) such that \( P(x) \) is true.

For example, the following is a formalization of the true statement that there exists a natural number \( n \) larger than 100 that is not divisible by 3:

\[
\exists n \in \mathbb{N} \ (n > 100) \land (\forall k \in \mathbb{N} \ k + k + k \neq n).
\]  

(1.10)

**For sufficiently large \( n \).** One expression that we will see come up time and again in this book is the claim that some statement \( P(n) \) is true “for sufficiently large \( n \)”. What this means is that there exists an integer \( N_0 \) such that \( P(n) \) is true for every \( n > N_0 \). We can formalize this as \( \exists N_0 \in \mathbb{N} \forall n > N_0 \) \( P(n) \).

### 1.4.6 Quantifiers for summations and products

The following shorthands for summing up or taking products of several numbers are often convenient. If \( S = \{s_0, \ldots, s_{n-1}\} \) is a finite set and \( f : S \to \mathbb{R} \) is a function, then we write \( \sum_{x \in S} f(x) \) as shorthand for

\[
f(s_0) + f(s_1) + f(s_2) + \ldots + f(s_{n-1}).
\]  

(1.11)

and \( \prod_{x \in S} f(x) \) as shorthand for

\[
f(s_0) \cdot f(s_1) \cdot f(s_2) \cdot \ldots \cdot f(s_{n-1}).
\]  

(1.12)

For example, the sum of the squares of all numbers from 1 to 100 can be written as

\[
\sum_{i=1}^{100} i^2.
\]  

(1.13)

Since summing up over intervals of integers is so common, there is a special notation for it. For every two integers, \( a \leq b \), \( \sum_{i=a}^{b} f(i) \) denotes \( \sum_{x \in S} f(i) \) where \( S = \{x \in \mathbb{Z} : a \leq x \leq b\} \). Hence, we can write the sum (1.13) as

\[
\sum_{i=1}^{100} i^2.
\]  

(1.14)

### 1.4.7 Parsing formulas: bound and free variables

In mathematics, as in coding, we often have symbolic “variables” or “parameters”. It is important to be able to understand, given some formula, whether a given variable is bound or free in this formula. For
example, in the following statement $n$ is free but $a$ and $b$ are bound by the $\exists$ quantifier:

$$\exists_{a,b \in \mathbb{N}} (a \neq 1) \land (a \neq n) \land (n = a \times b) \quad (1.15)$$

Since $n$ is free, it can be set to any value, and the truth of the statement (1.15) depends on the value of $n$. For example, if $n = 8$ then (1.15) is true, but for $n = 11$ it is false. (Can you see why?)

The same issue appears when parsing code. For example, in the following snippet from the C programming language

```c
for (int i=0 ; i<n ; i=i+1) {
    printf("*"n);
}
```

the variable $i$ is bound within the for block but the variable $n$ is free.

The main property of bound variables is that we can rename them (as long as the new name doesn’t conflict with another used variable) without changing the meaning of the statement. Thus for example the statement

$$\exists_{x,y \in \mathbb{N}} (x \neq 1) \land (x \neq n) \land (n = x \times y) \quad (1.16)$$

is equivalent to (1.15) in the sense that it is true for exactly the same set of $n$’s.

Similarly, the code

```c
for (int j=0 ; j<n ; j=j+1) {
    printf("*"n);
}
```

produces the same result as the code above that used $i$ instead of $j$.

**Remark 1.14 — Aside: mathematical vs programming notation.** Mathematical notation has a lot of similarities with programming language, and for the same reasons. Both are formalisms meant to convey complex concepts in a precise way. However, there are some cultural differences. In programming languages, we often try to use meaningful variable names such as `NumberOfVertices` while in math we often use short identifiers such as $n$. Part of it might have to do with the tradition of mathematical proofs as being handwritten and verbally presented, as opposed to typed up and compiled. Another reason is if the wrong variable name is used in a proof, at worst is causes confusion to readers; when the wrong variable name
is used in a program, planes might crash, patients might die, and rockets could explode.

One consequence of that is that in mathematics we often end up reusing identifiers, and also “run out” of letters and hence use Greek letters too, as well as distinguish between small and capital letters and different font faces. Similarly, mathematical notation tends to use quite a lot of “overloading”, using operators such as $+$ for a great variety of objects (e.g., real numbers, matrices, finite field elements, etc.), and assuming that the meaning can be inferred from the context.

Both fields have a notion of “types”, and in math we often try to reserve certain letters for variables of a particular type. For example, variables such as $i, j, k, l, m, n$ will often denote integers, and $\epsilon$ will often denote a small positive real number (see Section 1.7 for more on these conventions). When reading or writing mathematical texts, we usually don’t have the advantage of a “compiler” that will check type safety for us. Hence it is important to keep track of the type of each variable, and see that the operations that are performed on it “make sense”.

Kun’s book [Kun18] contains an extensive discussion on the similarities and differences between the cultures of mathematics and programming.

1.4.8 Asymptotics and Big-$O$ notation

“log log log $n$ has been proved to go to infinity, but has never been observed to do so.”, Anonymous, quoted by Carl Pomerance (2000)

It is often very cumbersome to describe precisely quantities such as running time and is also not needed, since we are typically mostly interested in the “higher order terms”. That is, we want to understand the scaling behavior of the quantity as the input variable grows. For example, as far as running time goes, the difference between an $n^5$-time algorithm and an $n^2$-time one is much more significant than the difference between an $100n^2 + 10n$ time algorithm and an $10n^2$ time algorithm. For this purpose, $O$-notation is extremely useful as a way to “declutter” our text and focus our attention on what really matters. For example, using $O$-notation, we can say that both $100n^2 + 10n$ and $10n^2$ are simply $\Theta(n^2)$ (which informally means “the same up to constant factors”), while $n^2 = o(n^5)$ (which informally means that $n^2$ is “much smaller than” $n^5$).

Generally (though still informally), if $F, G$ are two functions mapping natural numbers to non-negative reals, then “$F = O(G)$” means that $F(n) \leq G(n)$ if we don’t care about constant factors, while
“$F = o(G)$” means that $F$ is much smaller than $G$, in the sense that no matter by what constant factor we multiply $F$, if we take $n$ to be large enough then $G$ will be bigger (for this reason, sometimes $F = o(G)$ is written as $F \ll G$). We will write $F = \Theta(G)$ if $F = O(G)$ and $G = O(F)$, which one can think of as saying that $F$ is the same as $G$ if we don’t care about constant factors. More formally, we define Big-$O$ notation as follows:

**Definition 1.15 — Big-$O$ notation.** Let $\mathbb{R}_+ = \{x \in \mathbb{R} \mid x > 0\}$ be the set of positive real numbers. For two functions $F,G : \mathbb{N} \to \mathbb{R}_+$, we say that $F = O(G)$ if there exist numbers $a,N_0 \in \mathbb{N}$ such that $F(n) \leq a \cdot G(n)$ for every $n > N_0$. We say that $F = \Theta(G)$ if $F = O(G)$ and $G = O(F)$. We say that $F = \Omega(G)$ if $G = O(F)$.

We say that $F = o(G)$ if for every $\epsilon > 0$ there is some $N_0$ such that $F(n) < \epsilon G(n)$ for every $n > N_0$. We say that $F = \omega(G)$ if $G = o(F)$.

It’s often convenient to use “anonymous functions” in the context of $O$-notation. For example, when we write a statement such as $F(n) = O(n^3)$, we mean that $F = O(G)$ where $G$ is the function defined by $G(n) = n^3$. Chapter 7 in Jim Apsnes’ notes on discrete math provides a good summary of $O$ notation; see also this tutorial for a gentler and more programmer-oriented introduction.

$O$ is not equality. Using the equality sign for $O$-notation is extremely common, but is somewhat of a misnomer, since a statement such as $F = O(G)$ really means that $F$ is in the set $\{G' : \exists N,c$ s.t. $\forall n > N G'(n) \leq cG(n)\}$. If anything, it makes more sense to use inequalities and write $F \leq O(G)$ and $F \geq \Omega(G)$, reserving equality for $F = \Theta(G)$, and so we will sometimes use this notation too, but since the equality notation is quite firmly entrenched we often stick to it as well. (Some texts write $F \in O(G)$ instead of $F = O(G)$, but we will not use this notation.) Despite the misleading equality sign, you should remember that a statement such as $F = O(G)$ means that $F$ is “at most” $G$ in some rough sense when we ignore constants, and a statement such as $F = \Omega(G)$ means that $F$ is “at least” $G$ in the same rough sense.

1.4.9 Some “rules of thumb” for Big-$O$ notation

There are some simple heuristics that can help when trying to compare two functions $F$ and $G$:

- Multiplicative constants don’t matter in $O$-notation, and so if $F(n) = O(G(n))$ then $100F(n) = O(G(n))$.

- When adding two functions, we only care about the larger one. For example, for the purpose of $O$-notation, $n^3 + 100n^2$ is the same as
$n^3$, and in general in any polynomial, we only care about the larger exponent.

- For every two constants $a, b > 0$, $n^a = O(n^b)$ if and only if $a \leq b$, and $n^a = o(n^b)$ if and only if $a < b$. For example, combining the two observations above, $100n^2 + 10n + 100 = o(n^3)$.

- Polynomial is always smaller than exponential: $n^a = o(2^{n^\epsilon})$ for every two constants $a > 0$ and $\epsilon > 0$ even if $\epsilon$ is much smaller than $a$. For example, $100n^{100} = o(2^{\sqrt{n}})$.

- Similarly, logarithmic is always smaller than polynomial: $(\log n)^a$ (which we write as $\log^a n$) is $o(n^\epsilon)$ for every two constants $a, \epsilon > 0$. For example, combining the observations above, $100n^2 \log^{100} n = o(n^3)$.

---

**Remark 1.16 — Big O for other applications (optional).**

While Big-O notation is often used to analyze running time of algorithms, this is by no means the only application. We can use $O$ notation to bound asymptotic relations between any functions mapping integers to positive numbers. It can be used regardless of whether these functions are a measure of running time, memory usage, or any other quantity that may have nothing to do with computation. Here is one example which is unrelated to this book (and hence one that you can feel free to skip): one way to state the Riemann Hypothesis (one of the most famous open questions in mathematics) is that it corresponds to the conjecture that the number of primes between 0 and $n$ is equal to $\int_2^n \frac{1}{\ln x} \, dx$ up to an additive error of magnitude at most $O(\sqrt{n} \log n)$.

### 1.5 Proofs

Many people think of mathematical proofs as a sequence of logical deductions that starts from some axioms and ultimately arrives at a conclusion. In fact, some dictionaries define proofs that way. This is not entirely wrong, but at its essence mathematical proof of a statement $X$ is simply an argument that convinces the reader that $X$ is true beyond a shadow of a doubt.

To produce such a proof you need to:

1. Understand precisely what $X$ means.
2. Convince yourself that $X$ is true.
3. Write your reasoning down in plain, precise and concise English
   (using formulas or notation only when they help clarity).

   In many cases, the first part is the most important one. Understanding what a statement means is oftentimes more than halfway towards understanding why it is true. In third part, to convince the reader beyond a shadow of a doubt, we will often want to break down the reasoning to “basic steps”, where each basic step is simple enough to be “self evident”. The combination of all steps yields the desired statement.

1.5.1 Proofs and programs

There is a great deal of similarity between the process of writing proofs and that of writing programs, and both require a similar set of skills.

Writing a program involves:

1. Understanding what is the task we want the program to achieve.

2. Convincing yourself that the task can be achieved by a computer, perhaps by planning on a whiteboard or notepad how you will break it up to simpler tasks.

3. Converting this plan into code that a compiler or interpreter can understand, by breaking up each task into a sequence of the basic operations of some programming language.

   In programs as in proofs, step 1 is often the most important one. A key difference is that the reader for proofs is a human being and the reader for programs is a computer. (This difference is eroding with time as more proofs are being written in a machine verifiable form; moreover, to ensure correctness and maintainability of programs, it is important that they can be read and understood by humans.) Thus our emphasis is on readability and having a clear logical flow for our proof (which is not a bad idea for programs as well). When writing a proof, you should think of your audience as an intelligent but highly skeptical and somewhat petty reader, that will “call foul” at every step that is not well justified.

1.5.2 Proof writing style

A mathematical proof is a piece of writing, but it is a specific genre of writing with certain conventions and preferred styles. As in any writing, practice makes perfect, and it is also important to revise your drafts for clarity.

In a proof for the statement $X$, all the text between the words “Proof:” and “QED” should be focused on establishing that $X$ is true. Digressions, examples, or ruminations should be kept outside these
two words, so they do not confuse the reader. The proof should have a clear logical flow in the sense that every sentence or equation in it should have some purpose and it should be crystal-clear to the reader what this purpose is. When you write a proof, for every equation or sentence you include, ask yourself:

1. Is this sentence or equation stating that some statement is true?

2. If so, does this statement follow from the previous steps, or are we going to establish it in the next step?

3. What is the role of this sentence or equation? Is it one step towards proving the original statement, or is it a step towards proving some intermediate claim that you have stated before?

4. Finally, would the answers to questions 1-3 be clear to the reader? If not, then you should reorder, rephrase or add explanations.

Some helpful resources on mathematical writing include this handout by Lee, this handout by Hutching, as well as several of the excellent handouts in Stanford’s CS 103 class.

1.5.3 Patterns in proofs

“\text{If it was so, it might be; and if it were so, it would be; but as it isn’t, it ain’t. That’s logic.}”, Lewis Carroll, \textit{Through the looking-glass}.

Just like in programming, there are several common patterns of proofs that occur time and again. Here are some examples:

\textbf{Proofs by contradiction:} One way to prove that $X$ is true is to show that if $X$ was false it would result in a contradiction. Such proofs often start with a sentence such as “Suppose, towards a contradiction, that $X$ is false” and end with deriving some contradiction (such as a violation of one of the assumptions in the theorem statement). Here is an example:

\textbf{Lemma 1.17} There are no natural numbers $a, b$ such that $\sqrt{2} = \frac{a}{b}$.

\textit{Proof}. Suppose, towards a contradiction that this is false, and so let $a \in \mathbb{N}$ be the smallest number such that there exists some $b \in \mathbb{N}$ satisfying $\sqrt{2} = \frac{a}{b}$. Squaring this equation we get that $2 = \frac{a^2}{b^2}$ or $a^2 = 2b^2 \,(\ast)$. But this means that $a^2$ is even, and since the product of two odd numbers is odd, it means that $a$ is even as well, or in other words, $a = 2a'$ for some $a' \in \mathbb{N}$. Yet plugging this into $(\ast)$ shows that $4a'^2 = 2b^2$ which means $b^2 = 2a'^2$ is an even number as well. By the same considerations as above we get that $b$ is even and hence $a/2$ and
b/2 are two natural numbers satisfying $\frac{a}{2}/\frac{b}{2} = \sqrt{2}$, contradicting the minimality of $a$.

Proofs of a universal statement: Often we want to prove a statement $X$ of the form “Every object of type $O$ has property $P$.” Such proofs often start with a sentence such as “Let $o$ be an object of type $O$” and end by showing that $o$ has the property $P$. Here is a simple example:

**Lemma 1.18** For every natural number $n \in \mathbb{N}$, either $n$ or $n+1$ is even.

**Proof.** Let $n \in \mathbb{N}$ be some number. If $n/2$ is a whole number then we are done, since then $n = 2(n/2)$ and hence it is even. Otherwise, $n/2 + 1/2$ is a whole number, and hence $2(n/2 + 1/2) = n + 1$ is even.

Proofs of an implication: Another common case is that the statement $X$ has the form “$A$ implies $B$”. Such proofs often start with a sentence such as “Assume that $A$ is true” and end with a derivation of $B$ from $A$. Here is a simple example:

**Lemma 1.19** If $b^2 \geq 4ac$ then there is a solution to the quadratic equation $ax^2 + bx + c = 0$.

**Proof.** Suppose that $b^2 \geq 4ac$. Then $d = b^2 - 4ac$ is a non-negative number and hence it has a square root $s$. Thus $x = (-b + s)/(2a)$ satisfies

$$ax^2 + bx + c = a(-b + s)^2/(4a^2) + b(-b + s)/(2a) + c = (b^2 - 2bs + s^2)/(4a) + (-b^2 + bs)/(2a) + c.$$  \hspace{1cm} (1.17)

Rearranging the terms of (1.17) we get

$$s^2/(4a) + c - b^2/(4a) = (b^2 - 4ac)/(4a) + c - b^2/(4a) = 0 \hspace{1cm} (1.18)$$

Proofs of equivalence: If a statement has the form “$A$ if and only if $B$” (often shortened as “$A$ iff $B$”) then we need to prove both that $A$ implies $B$ and that $B$ implies $A$. We call the implication that $A$ implies $B$ the “only if” direction, and the implication that $B$ implies $A$ the “if” direction.

Proofs by combining intermediate claims: When a proof is more complex, it is often helpful to break it apart into several steps. That is, to prove the statement $X$, we might first prove statements $X_1$, $X_2$, and $X_3$ and then prove that $X_1 \wedge X_2 \wedge X_3$ implies $X$. (Recall that $\wedge$ denotes the logical AND operator.)
Proofs by case distinction: This is a special case of the above, where to prove a statement $X$ we split into several cases $C_1, \ldots, C_k$, and prove that (a) the cases are exhaustive, in the sense that one of the cases $C_i$ must happen and (b) go one by one and prove that each one of the cases $C_i$ implies the result $X$ that we are after.

Proofs by induction: We discuss induction and give an example in Section 1.6.1 below. We can think of such proofs as a variant of the above, where we have an unbounded number of intermediate claims $X_0, X_1, \ldots, X_k$, and we prove that $X_0$ is true, as well as that $X_0$ implies $X_1$, and that $X_0 \land X_1$ implies $X_2$, and so on and so forth. The website for CMU course 15-251 contains a useful handout on potential pitfalls when making proofs by induction.

Without loss of generality (w.l.o.g): This term can be initially quite confusing. It is essentially a way to simplify proofs by case distinctions. The idea is that if Case 1 is equal to Case 2 up to a change of variables or a similar transformation, then the proof of Case 1 will also imply the proof of Case 2. It is always a statement that should be viewed with suspicion. Whenever you see it in a proof, ask yourself if you understand why the assumption made is truly without loss of generality, and when you use it, try to see if the use is indeed justified. When writing a proof, sometimes it might be easiest to simply repeat the proof of the second case (adding a remark that the proof is very similar to the first one).

Remark 1.20 — Hierarchical Proofs (optional). Mathematical proofs are ultimately written in English prose. The well-known computer scientist Leslie Lamport argues that this is a problem, and proofs should be written in a more formal and rigorous way. In his manuscript he proposes an approach for structured hierarchical proofs, that have the following form:

- A proof for a statement of the form “If $A$ then $B$” is a sequence of numbered claims, starting with the assumption that $A$ is true, and ending with the claim that $B$ is true.
- Every claim is followed by a proof showing how it is derived from the previous assumptions or claims.
- The proof for each claim is itself a sequence of subclaims.

The advantage of Lamport’s format is that the role that every sentence in the proof plays is very clear. It is also much easier to transform such proofs into machine-checkable forms. The disadvantage is that
such proofs can be tedious to read and write, with less differentiation between the important parts of the arguments versus the more routine ones.

1.6 EXTENDED EXAMPLE: TOPOLOGICAL SORTING

In this section we will prove the following: every directed acyclic graph (DAG, see Definition 1.9) can be arranged in layers so that for all directed edges $u \rightarrow v$, the layer of $v$ is larger than the layer of $u$. This result is known as topological sorting and is used in many applications, including task scheduling, build systems, software package management, spreadsheet cell calculations, and many others (see Fig. 1.8). In fact, we will also use it ourselves later on in this book.

Figure 1.8: An example of topological sorting. We consider a directed graph corresponding to a prerequisite graph of the courses in some Computer Science program. The edge $u \rightarrow v$ means that the course $u$ is a prerequisite for the course $v$. A layering or “topological sorting” of this graph is the same as mapping the courses to semesters so that if we decide to take the course $v$ in semester $f(v)$, then we have already taken all the prerequisites for $v$ (i.e., its in-neighbors) in prior semesters.

We start with the following definition. A layering of a directed graph is a way to assign for every vertex $v$ a natural number (corresponding to its layer), such that $v$’s in-neighbors are in lower-numbered layers than $v$, and $v$’s out-neighbors are in higher-numbered layers. The formal definition is as follows:

Definition 1.21 — Layering of a DAG. Let $G = (V, E)$ be a directed graph. A layering of $G$ is a function $f : V \rightarrow \mathbb{N}$ such that for every edge $u \rightarrow v$ of $G$, $f(u) < f(v)$.

In this section we prove that a directed graph is acyclic if and only if it has a valid layering.

Theorem 1.22 — Topological Sort. Let $G$ be a directed graph. Then $G$ is acyclic if and only if there exists a layering $f$ of $G$.

To prove such a theorem, we need to first understand what it means. Since it is an “if and only if” statement, Theorem 1.22 corresponds to two statements:
Lemma 1.23  For every directed graph $G$, if $G$ is acyclic then it has a layering.

Lemma 1.24  For every directed graph $G$, if $G$ has a layering, then it is acyclic.

To prove Theorem 1.22 we need to prove both Lemma 1.23 and Lemma 1.24. Lemma 1.24 is actually not that hard to prove. Intuitively, if $G$ contains a cycle, then it cannot be the case that all edges on the cycle increase in layer number, since if we travel along the cycle at some point we must come back to the place we started from. The formal proof is as follows:

**Proof.** Let $G = (V, E)$ be a directed graph and let $f : V \rightarrow \mathbb{N}$ be a layering of $G$ as per Definition 1.21. Suppose, towards a contradiction, that $G$ is not acyclic, and hence there exists some cycle $u_0, u_1, \ldots, u_k$ such that $u_0 = u_k$ and for every $i \in [k]$ the edge $u_i \rightarrow u_{i+1}$ is present in $G$. Since $f$ is a layering, for every $i \in [k]$, $f(u_i) < f(u_{i+1})$, which means that

$$f(u_0) < f(u_1) < \cdots < f(u_k) \quad (1.19)$$

but this is a contradiction since $u_0 = u_k$ and hence $f(u_0) = f(u_k)$.

---

**Lemma 1.23** corresponds to the more difficult (and useful) direction. To prove it, we need to show how given an arbitrary DAG $G$, we can come up with a layering of the vertices of $G$ so that all edges “go up”.

---

1.6.1 Mathematical induction

There are several ways to prove Lemma 1.23. One approach to do is to start by proving it for small graphs, such as graphs with 1, 2 or 3 vertices (see Fig. 1.9, for which we can check all the cases, and then try to extend the proof for larger graphs. The technical term for this proof approach is *proof by induction*.

*Induction* is simply an application of the self-evident Modus Ponens rule that says that if...

---

**Figure 1.9:** Some examples of DAGs of one, two and three vertices, and valid ways to assign layers to the vertices.
(a) \( P \) is true

and

(b) \( P \) implies \( Q \)

then \( Q \) is true.

In the setting of proofs by induction we typically have a statement \( Q(k) \) that is parameterized by some integer \( k \), and we prove that (a) \( Q(0) \) is true, and (b) For every \( k > 0 \), if \( Q(0), \ldots, Q(k-1) \) are all true then \( Q(k) \) is true. (Usually proving (b) is the hard part, though there are examples where the “base case” (a) is quite subtle.) By applying Modus Ponens, we can deduce from (a) and (b) that \( Q(1) \) is true.

Once we did so, since we now know that both \( Q(0) \) and \( Q(1) \) are true, then we can use this and (b) to deduce (again using Modus Ponens) that \( Q(2) \) is true. We can repeat the same reasoning again and again to obtain that \( Q(k) \) is true for every \( k \). The statement (a) is called the “base case”, while (b) is called the “inductive step”. The assumption in (b) that \( Q(i) \) holds for \( i < k \) is called the “inductive hypothesis”.

(The form of induction described here is sometimes called “strong induction” as opposed to “weak induction” where we replace (b) by the statement (b’) that if \( Q(k-1) \) is true then \( Q(k) \) is true; weak induction can be thought of as the special case of strong induction where we don’t use the assumption that \( Q(0), \ldots, Q(k-2) \) are true.)

**Remark 1.25 — Induction and recursion.** Proofs by inductions are closely related to algorithms by recursion. In both cases we reduce solving a larger problem to solving a smaller instance of itself. In a recursive algorithm to solve some problem \( P \) on an input of length \( k \) we ask ourselves “what if someone handed me a way to solve \( P \) on instances smaller than \( k \)?”. In an inductive proof to prove a statement \( Q \) parameterized by a number \( k \), we ask ourselves “what if I already knew that \( Q(k’) \) is true for \( k’ < k \)?”. Both induction and recursion are crucial concepts for this course and Computer Science at large (and even other areas of inquiry, including not just mathematics but other sciences as well). Both can be confusing at first, but with time and practice they become clearer. For more on proofs by induction and recursion, you might find the following Stanford CS 103 handout, this MIT 6.00 lecture or this excerpt of the Lehman-Leighton book useful.

**1.6.2 Proving the result by induction**

There are several ways to use induction to prove Lemma 1.23 by induction. We will use induction on the number \( n \) of vertices, and so we will define the statement \( Q(n) \) as follows:
Q(n) is “For every DAG 𝐺 = (𝑉, 𝐸) with 𝑛 vertices, there is a layering of 𝐺.”

The statement for 𝑄(0) (where the graph contains no vertices) is trivial. Thus it will suffice to prove the following: for every 𝑛 > 0, if 𝑄(𝑛 − 1) is true then 𝑄(𝑛) is true.

To do so, we need to somehow find a way, given a graph 𝐺 of 𝑛 vertices, to reduce the task of finding a layering for 𝐺 into the task of finding a layering for some other graph 𝐺′ of 𝑛 − 1 vertices. The idea is that we will find a source of 𝐺: a vertex 𝑣 that has no in-neighbors. We can then assign to 𝑣 the layer 0, and layer the remaining vertices using the inductive hypothesis in layers 1, 2, ....

The above is the intuition behind the proof of Lemma 1.23, but when writing the formal proof below, we use the benefit of hindsight, and try to streamline what was a messy journey into a linear and easy-to-follow flow of logic that starts with the word “Proof” and ends with “QED” or the symbol ■. Discussions, examples and digressions can be very insightful, but we keep them outside the space delimited between these two words, where (as described by this excellent handout) “every sentence must be load bearing”. Just like we do in programming, we can break the proof into little “subroutines” or “functions” (known as lemmas or claims in math language), which will be smaller statements that help us prove the main result. However, the proof should be structured in a way that ensures that it is always crystal-clear to the reader in what stage we are of the proof. The reader should be able to tell what is the role of every sentence in the proof and which part it belongs to. We now present the formal proof of Lemma 1.23.

Proof of Lemma 1.23. Let 𝐺 = (𝑉, 𝐸) be a DAG and 𝑛 = |𝑉| be the number of its vertices. We prove the lemma by induction on 𝑛. The base case is 𝑛 = 0 where there are no vertices, and so the statement is trivially true. For the case of 𝑛 > 0, we make the inductive hypothesis that every DAG 𝐺′ of at most 𝑛 − 1 vertices has a layering.

We make the following claim:

Claim: 𝐺 must contain a vertex 𝑣 of in-degree zero.

Proof of Claim: Suppose otherwise that every vertex 𝑣 ∈ 𝑉 has an in-neighbor. Let 𝑣0 be some vertex of 𝐺, let 𝑣1 be an in-neighbor of 𝑣0, 𝑣2 be an in-neighbor of 𝑣1, and continue in this way for 𝑛 steps until we construct a list 𝑣0, 𝑣1, ..., 𝑣𝑛 such that for every 𝑖 ∈ [𝑛], 𝑣𝑖+1 is an in-neighbor of 𝑣𝑖, or in other words the edge 𝑣𝑖+1 → 𝑣𝑖 is present in the graph. Since there are only 𝑛 vertices in this graph, one of the 𝑛 + 1 vertices in this sequence must repeat itself, and so there exists 𝑖 < 𝑗 such that 𝑣𝑖 = 𝑣𝑗. But then the sequence 𝑣𝑗 → 𝑣𝑗−1 → ⋯ → 𝑣𝑖 is a cycle in 𝐺, contradicting our assumption that it is acyclic. (QED Claim)
Given the claim, we can let \( v_0 \) be some vertex of in-degree zero in \( G \), and let \( G' \) be the graph obtained by removing \( v_0 \) from \( G \). \( G' \) has \( n - 1 \) vertices and hence per the inductive hypothesis has a layering \( f' : (V \setminus \{v_0\}) \to \mathbb{N} \). We define \( f : V \to \mathbb{N} \) as follows:

\[
  f(v) = \begin{cases} 
    f'(v) + 1 & v \neq v_0 \\
    0 & v = v_0 
  \end{cases}. 
\]

We claim that \( f \) is a valid layering, namely that for every edge \( u \to v \), \( f(u) < f(v) \). To prove this, we split into cases:

- **Case 1:** \( u \neq v_0, v \neq v_0 \). In this case the edge \( u \to v \) exists in the graph \( G' \) and hence by the inductive hypothesis \( f'(u) < f'(v) \) which implies that \( f'(u) + 1 < f'(v) + 1 \).

- **Case 2:** \( u = v_0, v \neq v_0 \). In this case \( f(u) = 0 \) and \( f(v) = f'(v) + 1 > 0 \).

- **Case 3:** \( u \neq v_0, v = v_0 \). This case can’t happen since \( v_0 \) does not have in-neighbors.

- **Case 4:** \( u = v_0, v = v_0 \). This case again can’t happen since it means that \( v_0 \) is its own-neighbor — it is involved in a self loop which is a form cycle that is disallowed in an acyclic graph.

Thus, \( f \) is a valid layering for \( G \) which completes the proof.

---

**Reading a proof is no less of an important skill than producing one.** In fact, just like understanding code, it is a highly non-trivial skill in itself. Therefore I strongly suggest that you re-read the above proof, asking yourself at every sentence whether the assumption it makes are justified, and whether this sentence truly demonstrates what it purports to achieve. Another good habit is to ask yourself when reading a proof for every variable you encounter (such as \( u, i, G', f' \), etc. in the above proof) the following questions: (1) What type of variable is it? Is it a number? a graph? a vertex? a function? and (2) What do we know about it? Is it an arbitrary member of the set? Have we shown some facts about it?, and (3) What are we trying to show about it?

---

1.6.3 Minimality and uniqueness

**Theorem 1.22** guarantees that for every DAG \( G = (V, E) \) there exists some layering \( f : V \to \mathbb{N} \) but this layering is not necessarily *unique*. 
For example, if \( f : V \to \mathbb{N} \) is a valid layering of the graph then so is the function \( f' \) defined as \( f'(v) = 2 \cdot f(v) \). However, it turns out that the minimal layering is unique. A minimal layering is one where every vertex is given the smallest layer number possible. We now formally define minimality and state the uniqueness theorem:

**Theorem 1.26 — Minimal layering is unique.** Let \( G = (V, E) \) be a DAG. We say that a layering \( f : V \to \mathbb{N} \) is minimal if for every vertex \( v \in V \), if \( v \) has no in-neighbors then \( f(v) = 0 \) and if \( v \) has in-neighbors then there exists an in-neighbor \( u \) of \( v \) such that \( f(u) = f(v) - 1 \).

For every layering \( f, g : V \to \mathbb{N} \) of \( G \), if both \( f \) and \( g \) are minimal then \( f = g \).

The definition of minimality in Theorem 1.26 implies that for every vertex \( v \in V \), we cannot move it to a lower layer without making the layering invalid. If \( v \) is a source (i.e., has in-degree zero) then a minimal layering \( f \) must put it in layer \( 0 \), and for every other \( v \), if \( f(v) = i \), then we cannot modify this to set \( f(v) \leq i - 1 \) since there is an-neighbor \( u \) of \( v \) satisfying \( f(u) = i - 1 \). What Theorem 1.26 says is that a minimal layering \( f \) is unique in the sense that every other minimal layering is equal to \( f \).

**Proof Idea:**

The idea is to prove the theorem by induction on the layers. If \( f \) and \( g \) are minimal then they must agree on the source vertices, since both \( f \) and \( g \) should assign these vertices to layer \( 0 \). We can then show that if \( f \) and \( g \) agree up to layer \( i - 1 \), then the minimality property implies that they need to agree in layer \( i \) as well. In the actual proof we use a small trick to save on writing. Rather than proving the statement that \( f = g \) (or in other words that \( f(v) = g(v) \) for every \( v \in V \)), we prove the weaker statement that \( f(v) \leq g(v) \) for every \( v \in V \). (This is a weaker statement since the condition that \( f(v) \) is lesser or equal than to \( g(v) \) is implied by the condition that \( f(v) \) is equal to \( g(v) \).) However, since \( f \) and \( g \) are just labels we give to two minimal layerings, by simply changing the names “\( f \)” and “\( g \)” the same proof also shows that \( g(v) \leq f(v) \) for every \( v \in V \) and hence that \( f = g \).

\[ \star \]

**Proof of Theorem 1.26.** Let \( G = (V, E) \) be a DAG and \( f, g : V \to \mathbb{N} \) be two minimal valid layering of \( G \). We will prove that for every \( v \in V \), \( f(v) \leq g(v) \). Since we didn’t assume anything about \( f, g \) except their minimality, the same proof will imply that for every \( v \in V \), \( g(v) \leq f(v) \) and hence that \( f(v) = g(v) \) for every \( v \in V \), which is what we needed to show.
We will prove that $f(v) \leq g(v)$ for every $v \in V$ by induction on $i = f(v)$. The case $i = 0$ is immediate: since in this case $f(v) = 0$, $g(v)$ must be at least $f(v)$. For the case $i > 0$, by the minimality of $f$, if $f(v) = i$ then there must exist some in-neighbor $u$ of $v$ such that $f(u) = i - 1$. By the induction hypothesis we get that $g(u) \geq i - 1$, and since $g$ is a valid layering it must hold that $g(v) > g(u)$ which means that $g(v) \geq i = f(v)$.

The proof of Theorem 1.26 is fully rigorous, but is written in a somewhat terse manner. Make sure that you read through it and understand why this is indeed an airtight proof of the Theorem’s statement.

### 1.7 THIS BOOK: NOTATION AND CONVENTIONS

Most of the notation we use in this book is standard and is used in most mathematical texts. The main points where we diverge are:

- We index the natural numbers $\mathbb{N}$ starting with 0 (though many other texts, especially in computer science, do the same).

- We also index the set $[n]$ starting with 0, and hence define it as $\{0, \ldots, n-1\}$. In other texts it is often defined as $\{1, \ldots, n\}$. Similarly, we index our strings starting with 0, and hence a string $x \in \{0, 1\}^n$ is written as $x_0 x_1 \cdots x_{n-1}$.

- If $n$ is a natural number then $1^n$ does not equal the number 1 but rather this is the length $n$ string $11 \cdots 1$ (that is a string of $n$ ones). Similarly, $0^n$ refers to the length $n$ string $00 \cdots 0$.

- Partial functions are functions that are not necessarily defined on all inputs. When we write $f : A \to B$ this means that $f$ is a total function unless we say otherwise. When we want to emphasize that $f$ can be a partial function, we will sometimes write $f : A \to_p B$.

- As we will see later on in the course, we will mostly describe our computational problems in the terms of computing a Boolean function $f : \{0, 1\}^* \to \{0, 1\}$. In contrast, many other textbooks refer to the same task as deciding a language $L \subseteq \{0, 1\}^*$. These two viewpoints are equivalent, since for every set $L \subseteq \{0, 1\}^*$ there is a corresponding function $F$ such that $F(x) = 1$ if and only if $x \in L$. Computing partial functions corresponds to the task known in the literature as a solving a promise problem. Because the language notation is so prevalent in other textbooks, we will occasionally remind the reader of this correspondence.
• We use \[x\] and \([x]\) for the “ceiling” and “floor” operators that correspond to “rounding up” or “rounding down” a number to the nearest integer. We use \((x \mod y)\) to denote the “remainder” of \(x\) when divided by \(y\). That is, \((x \mod y) = x - y\lfloor x/y \rfloor\). In context when an integer is expected we’ll typically “silently round” the quantities to an integer. For example, if we say that \(x\) is a string of length \(\sqrt{n}\) then this means that \(x\) is of length \([\sqrt{n}]\). (We round up for the sake of convention, but in most such cases, it will not make a difference whether we round up or down.)

• Like most Computer Science texts, we default to the logarithm in base two. Thus, \(\log n\) is the same as \(\log_2 n\).

• We will also use the notation \(f(n) = \text{poly}(n)\) as a short hand for \(f(n) = n^{O(1)}\) (i.e., as shorthand for saying that there are some constants \(a, b\) such that \(f(n) \leq a \cdot n^b\) for every sufficiently large \(n\)). Similarly, we will use \(f(n) = \text{polylog}(n)\) as shorthand for \(f(n) = \text{poly}(\log n)\) (i.e., as shorthand for saying that there are some constants \(a, b\) such that \(f(n) \leq a \cdot (\log n)^b\) for every sufficiently large \(n\)).

• As in often the case in mathematical literature, we use the apostrophe character to enrich our set of identifiers. Typically if \(x\) denotes some object, then \(x', x'', \) etc. will denote other objects of the same type.

• To save on “cognitive load” we will often use round constants such as 10, 100, 1000 in the statements of both theorems and problem set questions. When you see such a “round” constant, you can typically assume that it has no special significance and was just chosen arbitrarily. For example, if you see a theorem of the form “Algorithm \(A\) takes at most \(1000 \cdot n^2\) steps to compute function \(F\) on inputs of length \(n\)” then probably the number 1000 is an arbitrary sufficiently large constant, and one could prove the same theorem with a bound of the form \(c \cdot n^2\) for a constant \(c\) that is smaller than 1000. Similarly, if a problem asks you to prove that some quantity is at least \(n/100\), it is quite possible that in truth the quantity is at least \(n/d\) for some constant \(d\) that is smaller than 100.

1.7.1 Variable name conventions
Like programming, mathematics is full of variables. Whenever you see a variable, it is always important to keep track of what is its type (e.g., whether the variable is a number, a string, a function, a graph, etc.). To make this easier, we try to stick to certain conventions and consistently use certain identifiers for variables of the same type. Some of these conventions are listed in Section 1.7.1 below. These conventions
are not immutable laws and we might occasionally deviate from them. Also, such conventions do not replace the need to explicitly declare for each new variable the type of object that it denotes.

<table>
<thead>
<tr>
<th>Identifier</th>
<th>Often denotes object of type</th>
</tr>
</thead>
<tbody>
<tr>
<td>$i, j, k, \ell, m, n$</td>
<td>Natural numbers (i.e., in $\mathbb{N} = {0, 1, 2, \ldots}$)</td>
</tr>
<tr>
<td>$e, \delta$</td>
<td>Small positive real numbers (very close to 0)</td>
</tr>
<tr>
<td>$x, y, z, w$</td>
<td>Typically strings in ${0, 1}^*$ though sometimes numbers or other objects. We often identify an object with its representation as a string.</td>
</tr>
<tr>
<td>$G$</td>
<td>A graph. The set of $G$’s vertices is typically denoted by $V$. Often $V = [n]$. The set of $G$’s edges is typically denoted by $E$.</td>
</tr>
<tr>
<td>$S$</td>
<td>Set</td>
</tr>
<tr>
<td>$f, g, h$</td>
<td>Functions. We often (though not always) use lowercase identifiers for finite functions, which map ${0, 1}^n$ to ${0, 1}^m$ (often $m = 1$).</td>
</tr>
<tr>
<td>$F, G, H$</td>
<td>Infinite (unbounded input) functions mapping ${0, 1}^<em>$ to ${0, 1}^</em>$ or ${0, 1}^*$ to ${0, 1}^m$ for some $m$. Based on context, the identifiers $G, H$ are sometimes used to denote functions and sometimes graphs.</td>
</tr>
<tr>
<td>$A, B, C$</td>
<td>Boolean circuits</td>
</tr>
<tr>
<td>$M, N$</td>
<td>Turing machines</td>
</tr>
<tr>
<td>$P, Q$</td>
<td>Programs</td>
</tr>
<tr>
<td>$T$</td>
<td>A function mapping $\mathbb{N}$ to $\mathbb{N}$ that corresponds to a time bound.</td>
</tr>
<tr>
<td>$c$</td>
<td>A positive number (often an unspecified constant; e.g., $T(n) = O(n)$ corresponds to the existence of $c$ s.t. $T(n) \leq c \cdot n$ every $n &gt; 0$). We sometimes use $a, b$ in a similar way.</td>
</tr>
<tr>
<td>$\Sigma$</td>
<td>Finite set (often used as the alphabet for a set of strings).</td>
</tr>
</tbody>
</table>

### 1.7.2 Some idioms

Mathematical texts often employ certain conventions or “idioms”. Some examples of such idioms that we use in this text include the following:

- “Let $X$ be ...”, “let $X$ denote ...”, or “let $X = ...$”: These are all different ways for us to say that we are defining the symbol $X$ to stand for whatever expression is in the .... When $X$ is a property of some objects we might define $X$ by writing something along the
lines of “We say that ... has the property $X$ if ....”. While we often try to define terms before they are used, sometimes a mathematical sentence reads easier if we use a term before defining it, in which case we add “Where $X$ is ...” to explain how $X$ is defined in the preceding expression.

- **Quantifiers**: Mathematical texts involve many quantifiers such as “for all” and “exists”. We sometimes spell these in words as in “for all $i \in \mathbb{N}$” or “there is $x \in \{0, 1\}$”, and sometimes use the formal symbols $\forall$ and $\exists$. It is important to keep track on which variable is quantified in what way the dependencies between the variables. For example, a sentence fragment such as “for every $k > 0$ there exists $n$” means that $n$ can be chosen in a way that depends on $k$. The order of quantifiers is important. For example, the following is a true statement: “for every natural number $k > 1$ there exists a prime number $n$ such that $n$ divides $k$.” In contrast, the following statement is false: “there exists a prime number $n$ such that for every natural number $k > 1$, $n$ divides $k$.”

- **Numbered equations, theorems, definitions**: To keep track of all the terms we define and statements we prove, we often assign them a (typically numeric) label, and then refer back to them in other parts of the text.

- (i.e.,), (e.g.,): Mathematical texts tend to contain quite a few of these expressions. We use $X$ (i.e., $Y$) in cases where $Y$ is equivalent to $X$ and $X$ (e.g., $Y$) in cases where $Y$ is an example of $X$ (e.g., one can use phrases such as “a natural number (i.e., a non-negative integer)” or “a natural number (e.g., 7)”).

- “Thus”, “Therefore”, “We get that”: This means that the following sentence is implied by the preceding one, as in “The $n$-vertex graph $G$ is connected. Therefore it contains at least $n - 1$ edges.” We sometimes use “indeed” to indicate that the following text justifies the claim that was made in the preceding sentence as in “The $n$-vertex graph $G$ has at least $n - 1$ edges. Indeed, this follows since $G$ is connected.”

- **Constants**: In Computer Science, we typically care about how our algorithms’ resource consumption (such as running time) scales with certain quantities (such as the length of the input). We refer to quantities that do not depend on the length of the input as constants and so often use statements such as “there exists a constant $c > 0$ such that for every $n \in \mathbb{N}$, Algorithm $A$ runs in at most $c \cdot n^2$ steps on inputs of length $n$.” The qualifier “constant” for $c$ is not strictly needed but is added to emphasize that $c$ here is a fixed number independent of $n$. 
In fact sometimes, to reduce cognitive load, we will simply replace $c$ by a sufficiently large round number such as 10, 100, or 1000, or use $O$-notation and write “Algorithm A runs in $O(n^2)$ time.”

Lecture Recap

- The basic “mathematical data structures” we’ll need are *numbers*, *sets*, *tuples*, *strings*, *graphs* and *functions*.
- We can use basic objects to define more complex notions. For example, *graphs* can be defined as a list of *pairs*.
- Given precise definitions of objects, we can state unambiguous and precise *statements*. We can then use mathematical *proofs* to determine whether these statements are true or false.
- A mathematical proof is not a formal ritual but rather a clear, precise and “bulletproof” argument certifying the truth of a certain statement.
- Big-$O$ notation is an extremely useful formalism to suppress less significant details and allow us to focus on the high level behavior of quantities of interest.
- The only way to get comfortable with mathematical notions is to apply them in the contexts of solving problems. You should expect to need to go back time and again to the definitions and notation in this chapter as you work through problems in this course.

1.8 EXERCISES

**Exercise 1.1 — Logical expressions.**

a. Write a logical expression $\varphi(x)$ involving the variables $x_0, x_1, x_2$ and the operators $\land$ (AND), $\lor$ (OR), and $\neg$ (NOT), such that $\varphi(x)$ is true if the majority of the inputs are *True*.

b. Write a logical expression $\varphi(x)$ involving the variables $x_0, x_1, x_2$ and the operators $\land$ (AND), $\lor$ (OR), and $\neg$ (NOT), such that $\varphi(x)$ is true if the sum $\sum_{i=0}^{2} x_i$ (identifying “true” with 1 and “false” with 0) is *odd*.

**Exercise 1.2 — Quantifiers.** Use the logical quantifiers $\forall$ (for all), $\exists$ (there exists), as well as $\land, \lor, \neg$ and the arithmetic operations $+, \times, =, >, <$ to write the following:

a. An expression $\varphi(n, k)$ such that for every natural numbers $n, k$, $\varphi(n, k)$ is true if and only if $k$ divides $n$. 
b. An expression \( \varphi(n) \) such that for every natural number \( n \), \( \varphi(n) \) is true if and only if \( n \) is a power of three.

**Exercise 1.3** Describe the following statement in English words:
\[
\forall n \in \mathbb{N} \exists p > n \forall a, b \in \mathbb{N} (a \times b \neq p) \lor (a = 1).
\]

**Exercise 1.4 — Set construction notation.** Describe in words the following sets:

a. \( S = \{ x \in \{0, 1\}^{100} : \forall i \in \{0, \ldots, 99\} x_i = x_{99-i} \} \)

b. \( T = \{ x \in \{0, 1\}^* : \forall i, j \in \{2, \ldots, |x| - 1\} i \cdot j \neq |x| \} \)

**Exercise 1.5 — Existence of one to one mappings.** For each one of the following pairs of sets \((S, T)\), prove or disprove the following statement: there is a one to one function \( f \) mapping \( S \) to \( T \).

a. Let \( n > 10 \). \( S = \{0, 1\}^n \) and \( T = [n] \times [n] \times [n] \).

b. Let \( n > 10 \). \( S \) is the set of all functions mapping \( \{0, 1\}^n \) to \( \{0, 1\} \).
   \( T = \{0, 1\}^{3n} \).

c. Let \( n > 100 \). \( S = \{ k \in [n] \mid k \text{ is prime} \} \), \( T = \{0, 1\}^{[\log n - 1]} \).

**Exercise 1.6 — Inclusion Exclusion.**

a. Let \( A, B \) be finite sets. Prove that
\[
|A \cup B| = |A| + |B| - |A \cap B|.
\]

b. Let \( A_0, \ldots, A_{k-1} \) be finite sets. Prove that
\[
|A_0 \cup \cdots \cup A_{k-1}| \geq \sum_{i=0}^{k-1} |A_i| - \sum_{0 \leq i < j < k} |A_i \cap A_j|.
\]

c. Let \( A_0, \ldots, A_{k-1} \) be finite subsets of \( \{1, \ldots, n\} \), such that \( |A_i| = m \) for every \( i \in [k] \). Prove that if \( k > 100n \), then there exist two distinct sets \( A_i, A_j \) s.t. \( |A_i \cap A_j| \geq m^2/(10n) \).

**Exercise 1.7** Prove that if \( S, T \) are finite and \( F : S \to T \) is one to one then \( |S| \leq |T| \).

**Exercise 1.8** Prove that if \( S, T \) are finite and \( F : S \to T \) is onto then \( |S| \geq |T| \).

**Exercise 1.9** Prove that for every finite \( S, T \), there are \((|T| + 1)^{|S|}\) partial functions from \( S \) to \( T \).
Exercise 1.10 Suppose that \( \{S_n\}_{n \in \mathbb{N}} \) is a sequence such that \( S_0 \leq 10 \) and for \( n > 1 \) \( S_n \leq 5S_{\lfloor \frac{n}{5} \rfloor} + 2n \). Prove by induction that \( S_n \leq 100n \log n \) for every \( n \).

Exercise 1.11 Prove that for every undirected graph \( G \) of 100 vertices, if every vertex has degree at most 4, then there exists a subset \( S \) of at 20 vertices such that no two vertices in \( S \) are neighbors of one another.

Exercise 1.12 — \( O \)-notation. For every pair of functions \( F, G \) below, determine which of the following relations holds: \( F = O(G) \), \( F = \Omega(G) \), \( F = o(G) \) or \( F = \omega(G) \).

a. \( F(n) = n, G(n) = 100n \).

b. \( F(n) = n, G(n) = \sqrt{n} \).

c. \( F(n) = n \log n, G(n) = 2^{(\log(n))^2} \).

d. \( F(n) = \sqrt{n}, G(n) = 2^{\log n} \).

e. \( F(n) = \left( \binom{n}{\lfloor 0.2n \rfloor} \right), G(n) = 2^{0.1n} \) (where \( \binom{n}{k} \) is the number of \( k \)-sized subsets of a set of size \( n \)) and \( g(n) = 2^{0.1n} \). See footnote for hint.\(^7\)

Exercise 1.13 Give an example of a pair of functions \( F, G : \mathbb{N} \to \mathbb{N} \) such that neither \( F = O(G) \) nor \( G = O(F) \) holds.

Exercise 1.14 Prove that for every undirected graph \( G \) on \( n \) vertices, if \( G \) has at least \( n \) edges then \( G \) contains a cycle.

Exercise 1.15 Prove that for every undirected graph \( G \) of 1000 vertices, if every vertex has degree at most 4, then there exists a subset \( S \) of at least 200 vertices such that no two vertices in \( S \) are neighbors of one another.

1.9 BIBLIOGRAPHICAL NOTES

The heading “A Mathematician’s Apology”, refers to Hardy’s classic book [Har41]. Even when Hardy is wrong, he is very much worth reading.

There are many online sources for the mathematical background needed for this book. In particular, the lecture notes for MIT 6.042 “Mathematics for Computer Science” [LLM18] are extremely comprehensive, and videos and assignments for this course are available.
online. Similarly, Berkeley CS 70: “Discrete Mathematics and Probability Theory” has extensive lecture notes online.

Other sources for discrete mathematics are Rosen [Ros19] and Jim Aspens’ online book [Asp18]. Lewis and Zax [LZ19], as well as the online book of Fleck [Fle18], give a more gentle overview of the much of the same material. Solow [Sol14] is a good introduction to proof reading and writing. Kun [Kun18] gives an introduction to mathematics aimed at readers with programming background. Stanford’s CS 103 course has a wonderful collections of handouts on mathematical proof techniques and discrete mathematics.

The word graph in the sense of Definition 1.3 was coined by the mathematician Sylvester in 1878 in analogy with the chemical graphs used to visualize molecules. There is an unfortunate confusion between this term and the more common usage of the word “graph” as a way to plot data, and in particular a plot of some function $f(x)$ as a function of $x$. One way to relate these two notions is to identify every function $f : A \rightarrow B$ with the directed graph $G_f$ over the vertex set $V = A \cup B$ such that $G_f$ contains the edge $x \rightarrow f(x)$ for every $x \in A$. In a graph $G_f$ constructed in this way, every vertex in $A$ has out-degree equal to one. If the function $f$ is one to one then every vertex in $B$ has in-degree at most one. If the function $f$ is onto then every vertex in $B$ has in-degree at least one. If $f$ is a bijection then every vertex in $B$ has in-degree exactly equal to one.

Carl Pomerance’s quote is taken from the home page of Doron Zeilberger.
2

Computation and Representation

To a first approximation, computation is a process that maps an input to an output.

When discussing computation, it is essential to separate the question of what is the task we need to perform (i.e., the specification) from the question of how we achieve this task (i.e., the implementation).

For example, as we’ve seen, there is more than one way to achieve the computational task of computing the product of two integers.
In this chapter we focus on the what part, namely defining computational tasks. For starters, we need to define the inputs and outputs. Capturing all the potential inputs and outputs that we might ever want to compute on seems challenging, since computation today is applied to a wide variety of objects. We do not compute merely on numbers, but also on texts, images, videos, connection graphs of social networks, MRI scans, gene data, and even other programs. We will represent all these objects as strings of zeroes and ones, that is objects such as $0011101$ or $1011$ or any other finite list of 1’s and 0’s. (This choice is for convenience: there is nothing “holy” about zeroes and ones, and we could have used any other finite collection of symbols.)

Today, we are so used to the notion of digital representation that we are not surprised by the existence of such an encoding. But it is actually a deep insight with significant implications. Many animals can convey a particular fear or desire, but what is unique about humans is language: we use a finite collection of basic symbols to describe a potentially unlimited range of experiences. Language allows transmission of information over both time and space and enables societies that span a great many people and accumulate a body of shared knowledge over time.

Over the last several decades, we have seen a revolution in what we can represent and convey in digital form. We can capture experiences with almost perfect fidelity, and disseminate it essentially instantaneously to an unlimited audience. Moreover, once information is in digital form, we can compute over it, and gain insights from data that were not accessible in prior times. At the heart of this revolution is the simple but profound observation that we can represent an unbounded variety of objects using a finite set of symbols (and in fact using only the two symbols 0 and 1).

In later chapters, we will typically take such representations for granted, and hence use expressions such as “program $P$ takes $x$ as input” when $x$ might be a number, a vector, a graph, or any other object, when we really mean that $P$ takes as input the representation of $x$ as a binary string. However, in this chapter we will dwell a bit more on how we can construct such representations.

2.1 DEFINING REPRESENTATIONS

Every time we store numbers, images, sounds, databases, or other objects on a computer, what we actually store in the computer’s memory is the representation of these objects. Moreover, the idea of representation is not restricted to digital computers. When we write down text or make a drawing we are representing ideas or experiences as sequences of symbols (which might as well be strings of zeroes and ones). Even
our brain does not store the actual sensory inputs we experience, but rather only a representation of them.

To use objects such as numbers, images, graphs, or others as inputs for computation, we need to define precisely how to represent these objects as binary strings. A representation scheme is a way to map an object \( x \) to a binary string \( E(x) \in \{0, 1\}^* \). For example, a representation scheme for natural numbers is a function \( E : \mathbb{N} \rightarrow \{0, 1\}^* \). Of course, we cannot merely represent all numbers as the string "0011" (for example). A minimal requirement is that if two numbers \( x \) and \( x' \) are different then they would be represented by different strings. Another way to say this is that we require the encoding function \( E \) to be one to one.

### 2.1.1 Representing natural numbers

We now show how we can represent natural numbers as binary strings. Over the years people have represented numbers in a variety of ways, including Roman numerals, tally marks, our own Hindu-Arabic decimal system, and many others. We can use any one of those as well as many others to represent a number as a string (see Fig. 2.3). However, for the sake of concreteness, we use the binary basis as our default representation of natural numbers as strings. For example, we represent the number six as the string \( 110 \) since \( 1 \cdot 2^2 + 1 \cdot 2^1 + 0 \cdot 2^0 = 6 \), and similarly we represent the number thirty-five as the string \( y = 100011 \) which satisfies \( \sum_{i=0}^{5} y_i \cdot 2^{|y|−i−1} = 35 \).

Some more examples are given in the table below.

<table>
<thead>
<tr>
<th>Number (decimal representation)</th>
<th>Number (binary representation)</th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>0</td>
</tr>
<tr>
<td>1</td>
<td>1</td>
</tr>
<tr>
<td>2</td>
<td>10</td>
</tr>
<tr>
<td>5</td>
<td>101</td>
</tr>
<tr>
<td>16</td>
<td>10000</td>
</tr>
<tr>
<td>40</td>
<td>101000</td>
</tr>
<tr>
<td>53</td>
<td>110101</td>
</tr>
<tr>
<td>389</td>
<td>110000101</td>
</tr>
<tr>
<td>3750</td>
<td>111010100110</td>
</tr>
</tbody>
</table>

1 We could just as well used the reverse order and represent 35 as the string 110001. Similarly, we could have just as well used the empty string "" to represent 0. Such low level choices do not make a difference for us, though they do matter in computing practice, see the bibliographical notes Section 2.7.

Figure 2.3: Representing each one the digits 0, 1, 2, ..., 9 as a 12 × 8 bitmap image, which can be thought of as a string in \{0, 1\}^{96}. Using this scheme we can represent a natural number \( n \) of \( n \) decimal digits as a string in \{0, 1\}^{96n}. Image taken from blog post of A. C. Andersen.
The notation \(|n/10|\) corresponds to "chopping off" the least significant decimal digit (e.g., \(|457/10| = |45.7| = 45\), the number \(|n/2|\) corresponds to the "chopping off" the least significant binary digit. Hence the binary representation can be formally defined as the following function \(NtS : \mathbb{N} \rightarrow \{0, 1\}^*\) (\(NtS\) stands for "natural numbers to strings"):

\[
NtS(n) = \begin{cases}
0 & n = 0 \\
1 & n = 1 \\
NtS(|n/2|) \text{parity}(n) & n > 1
\end{cases}
\]  

(2.1)

where \(\text{parity} : \mathbb{N} \rightarrow \{0, 1\}\) is the function defined as \(\text{parity}(n) = 0\) if \(n\) is even and \(\text{parity}(n) = 1\) if \(n\) is odd, and as usual, for strings \(x, y \in \{0, 1\}^*\), \(xy\) denotes the concatenation of \(x\) and \(y\). The function \(NtS\) is defined recursively: for every \(n > 0\) we define \(\text{rep}(n)\) in terms of the representation of the smaller number \(|n/2|\). It is also possible to define \(NtS\) non-recursively, see Exercise 2.2.

Throughout most of this book, the particular choices of representation of numbers as binary strings would not matter much, we just need to know that such a representation exists. In fact, for most purposes, we can even use the simpler representation of mapping a natural number \(n\) to the length-\(n\) all-zero string \(0^n\).

**Remark 2.1** — Binary representation in python (optional).
We can implement the binary representation in Python as follows:

```python
def NtS(n):  # natural numbers to strings
    if n > 1:
        return NtS(n // 2) + str(n % 2)
    else:
        return str(n % 2)

print(NtS(236))  # 11101100
print(NtS(19))   # 10011
```

We can also use Python to implement the inverse transformation, mapping a string back to the natural number it represents.

```python
def StN(x):  # String to number
    k = len(x)-1
    return sum(int(x[i])*(2**i) for i in range(k+1))

print(StN(NtS(236)))  # 236
```

While the Babylonians already invented a positional system much earlier, the decimal positional system we use today was invented by Indian mathematicians around the third century. It was taken up by Arab mathematicians in the 8th century. It was mainly introduced to Europe in the 1202 book “Liber Abaci” by Leonardo of Pisa, also known as Fibonacci, but did not displace Roman numerals in common usage until the 15th century.

Remark 2.2 — Programming examples. In this book, we sometimes use code examples as in Remark 2.1. The point is always to emphasize that certain computations can be achieved concretely, rather than illustrating the features of Python or any other programming language. Indeed, one of the messages of this book is that all programming languages are in a certain precise sense equivalent to one another, and hence we could have just as well used JavaScript, C, COBOL, Visual Basic or even BrainF*ck. This book is not about programming, and it is absolutely OK if you are not familiar with Python or do not follow code examples such as those in Remark 2.1.

2.1.2 Meaning of representations (discussion)

It is natural for us to think of 236 as the “actual” number, and of 11101100 as “merely” its representation. However, for most Europeans in the middle ages CCXXXVI would be the “actual” number and 236 (if they have even heard about it) would be the weird Hindu-Arabic positional representation.2 When our AI robot overlords materialize, they will probably think of 11101100 as the “actual” number and of 236 as “merely” a representation that they need to use when they give commands to humans.

So what is the “actual” number? This is a question that philosophers of mathematics have pondered over throughout history. Plato argued that mathematical objects exist in some ideal sphere of existence (that to a certain extent is more “real” than the world we perceive via our senses, as this latter world is merely the shadow of this ideal sphere). In Plato’s vision, the symbols 236 are merely notation for some ideal object, that, in homage to the late musician, we can refer to as “the number commonly represented by 236”.

The Austrian philosopher Ludwig Wittgenstein, on the other hand, argued that mathematical objects do not exist at all, and the only things that exist are the actual marks on paper that make up 236, 11101100 or CCXXXVI. In Wittgenstein’s view, mathematics is merely about formal manipulation of symbols that do not have any inherent meaning. You can think of the “actual” number as (somewhat recursively) “that thing which is common to 236, 11101100 and CCXXXVI and all other past and future representations that are meant to capture the same object”.

While reading this book, you are free to choose your own philosophy of mathematics, as long as you maintain the distinction between the mathematical objects themselves and the various particular
choices of representing them, whether as splotches of ink, pixels on a screen, zeroes and one, or any other form.

2.2 REPRESENTATIONS BEYOND NATURAL NUMBERS

We have seen that natural numbers can be represented as binary strings. We now show that the same is true for other types of objects, including (potentially negative) integers, rational numbers, vectors, lists, graphs and many others. In many instances, choosing the “right” string representation for a piece of data is highly nontrivial, and finding the “best” one (e.g., most compact, best fidelity, most efficiently manipulable, robust to errors, most informative features, etc.) is the object of intense research. But for now, we focus on presenting some simple representations for various objects that we would like to use as inputs and outputs for computation.

2.2.1 Representing (potentially negative) integers

Since we can represent natural numbers as strings, we can represent the full set of integers (i.e., members of the set \( \mathbb{Z} = \{\ldots, -3, -2, -1, 0, +1, +2, +3, \ldots\} \)) by adding one more bit that represents the sign. To represent a (potentially negative) number \( m \), we prepend to the representation of the natural number \(|m|\) a bit \( \sigma \) that equals 0 if \( m \geq 0 \) and equals 1 if \( m < 0 \). Formally, we define the function \( ZtS : \mathbb{Z} \rightarrow \{0, 1\}^* \) as follows

\[
ZtS(m) = \begin{cases} 
0 & \text{if } m \geq 0 \\
1 & \text{if } m < 0 
\end{cases}
\]

where \( NtS \) is defined as in (2.1).

While the encoding function of a representation needs to be one to one, it does not have to be onto. For example, in the representation above there is no number that is represented by the empty string but it is still a fine representation, since every integer is represented uniquely by some string.

Remark 2.3 — Interpretation and context. Given a string \( y \in \{0, 1\}^* \), how do we know if it’s “supposed” to represent a (nonnegative) natural number or a (potentially negative) integer? For that matter, even if we know \( y \) is “supposed” to be an integer, how do we know what representation scheme it uses? The short answer is that we do not necessarily know this information, unless it is supplied from the context. (In programming languages, the compiler or interpreter determines the representation of the sequence of bits corresponding to a variable based on the variable’s type.) We can treat the same string \( y \) as representing a
natural number, an integer, a piece of text, an image, or a green gremlin. Whenever we say a sentence such as “let \( n \) be the number represented by the string \( y \),” we will assume that we are fixing some canonical representation scheme such as the ones above. The choice of the particular representation scheme will rarely matter, except that we want to make sure to stick with the same one for consistency.

2.2.2 Two’s complement representation (optional)

Section 2.2.1’s approach of representing an integer using a specific “sign bit” is known as the Signed Magnitude Representation and was used in some early computers. However, the two’s complement representation is much more common in practice. The two’s complement representation of an integer \( k \) in the set \( \{-2^n, -2^n+1, \ldots, 2^n-1\} \) is the string \( Z_{TS_n}(k) \) of length \( n + 1 \) defined as follows:

\[
Z_{TS_n}(k) = \begin{cases} 
Nts_{n+1}(k) & 0 \leq k \leq 2^n - 1 \\
Nts_{n+1}(2^{n+1} + k) & -2^n \leq k \leq -1 
\end{cases},
\]

(2.3)

where \( Nts_{\ell}(m) \) denotes the standard binary representation of a number \( m \in \{0, \ldots, 2^\ell\} \) as string of length \( \ell \), padded with leading zeros as needed. For example, if \( n = 3 \) then \( Z_{TS_3}(1) = Nts_4(1) = 0001 \), \( Z_{TS_3}(2) = Nts_4(2) = 0010 \), \( Z_{TS_3}(-1) = Nts_4(16 - 1) = 1111 \), and \( Z_{TS_3}(-8) = Nts_4(16 - 8) = 1000 \). If \( k \) is a negative number larger or equal to \( -2^n \) then \( 2^{n+1} + k \) is a number between \( 2^n \) and \( 2^{n+1} - 1 \).

Hence the two’s complement representation of such a number \( k \) is a string of length \( n + 1 \) with its first digit equal to 1.

Another way to say this is that we represent a potentially negative number \( k \in \{-2^n, \ldots, 2^n - 1\} \) as the non-negative number \( k \mod 2^{n+1} \) (see also Fig. 2.4). This means that if two (potentially negative) numbers \( k \) and \( k' \) are not too large (i.e., \( |k| + |k'| < 2^{n+1} \)), then we can compute the representation of \( k + k' \) by adding modulo \( 2^{n+1} \) the representations of \( k \) and \( k' \) as if they were non-negative integers. This property of the two’s complement representation is its main attraction since, depending on their architectures, microprocessors can often perform arithmetic operations modulo \( 2^n \) very efficiently (for certain values of \( w \) such as 32 and 64). Many systems leave it to the programmer to check that values are not too large and will carry out this modular arithmetic regardless of the size of the numbers involved. For this reason, in some systems adding two large positive numbers can result in a negative number (e.g., adding \( 2^n - 100 \) and \( 2^n - 200 \) might result in \(-300 \) since \(-300 \mod 2^{n+1} = 2^{n+1} - 300 \), see also Fig. 2.4).

![Figure 2.4](image-url)
2.2.3 Rational numbers, and representing pairs of strings

We can represent a rational number of the form $\frac{a}{b}$ by representing the two numbers $a$ and $b$. However, merely concatenating the representations of $a$ and $b$ will not work. For example, the binary representation of 4 is 100 and the binary representation of 43 is 101011, but the concatenation 10010111 of these strings is also the concatenation of the representation 10010 of 18 and the representation 1011 of 11. Hence, if we used such simple concatenation then we would not be able to tell if the string 10010111 is supposed to represent $\frac{4}{43}$ or $\frac{18}{11}$.

We tackle this by giving a general representation for pairs of strings. If we were using a pen and paper, we would just use a separator symbol such as $\parallel$ to represent, for example, the pair consisting of the numbers represented by 10 and 110001 as the length-9 string “01110001”. In other words, there is a one to one map $F$ from pairs of strings $x, y \in \{0, 1\}^*$ into a single string $z$ over the alphabet $\Sigma = \{0, 1, \parallel\}$. Using such separators is similar to the way we use spaces and punctuation to separate words in English. By adding a little redundancy, we achieve the same effect in the digital domain. We can map the three-element set $\Sigma$ to the three-element set $\{\text{00}, \text{11}, \text{01}\} \subset \{0, 1\}^2$ in a one-to-one fashion, and hence encode a length $n$ string $z \in \Sigma^*$ as a length $2n$ string $w \in \{0, 1\}^*$.

Our final representation for rational numbers is obtained by composing the following steps:

1. Representing a non-negative rational number as a pair of natural numbers.

2. Representing a natural number by a string via the binary representation.

3. Combining 1 and 2 to obtain a representation of a rational number as a pair of strings.

4. Representing a pair of strings over $\{0, 1\}$ as a single string over $\Sigma = \{0, 1, \parallel\}$.

5. Representing a string over $\Sigma$ as a longer string over $\{0, 1\}$.

**Example 2.4 — Representing a rational number as a string.** Consider the rational number $r = -\frac{5}{8}$. We represent $-5$ as 1101 and $+8$ as 01000, and so we can represent $r$ as the pair of strings (1101, 01000) and represent this pair as the length 10 string 1101110000 over the alphabet $\{0, 1, \parallel\}$. Now, applying the map $0 \mapsto \text{00}$, $1 \mapsto \text{11}$,
|| → 01, we can represent the latter string as the length 20 string
\[ s = 111100110011000000 \]
over the alphabet \{0, 1\}.

The same idea can be used to represent triples of strings, quadruplets, and so on as a string. Indeed, this is one instance of a very general principle that we use time and again in both the theory and practice of computer science (for example, in Object Oriented programming):

**Big Idea 1** If we can represent objects of type \( T \) as strings, then we can represent tuples of objects of type \( T \) as strings as well.

Repeating the same idea, once we can represent objects of type \( T \), we can also represent lists of lists of such objects, and even lists of lists of lists and so on and so forth. We will come back to this point when we discuss **prefix free encoding** in **Section 2.4.2**.

### 2.3 Representing Real Numbers

The set of real numbers \( \mathbb{R} \) contains all numbers including positive, negative, and fractional, as well as irrational numbers such as \( \pi \) or \( e \). Every real number can be approximated by a rational number, and thus we can represent every real number \( x \) by a rational number \( a/b \) that is very close to \( x \). For example, we can represent \( \pi \) by \( 22/7 \) within an error of about \( 10^{-3} \). If we want a smaller error (e.g., about \( 10^{-4} \)) then we can use \( 311/99 \), and so on and so forth.

![Figure 2.5: The floating point representation of a real number \( x \in \mathbb{R} \) is its approximation as a number of the form \( \sigma b \cdot 2^e \) where \( \sigma \in \{±1\} \), \( e \) is an (potentially negative) integer, and \( b \) is a rational number between 1 and 2 expressed as a binary fraction \( 1.b_1 b_2 \ldots b_k \) for some \( b_1, \ldots, b_k \in \{0, 1\} \) (that is \( b = 1 + b_1/2 + b_2/4 + \ldots + b_k/2^k \)). Commonly-used floating point representations fix the numbers \( \ell \) and \( k \) of bits to represent \( e \) and \( b \) respectively. In the example above, assuming we use two’s complement representation for \( e \), the number represented is \( -1 \times 2^5 \times (1 + 1/2 + 1/4 + 1/64 + 1/512) = -56.5625 \).](image)

The above representation of real numbers via rational numbers that approximate them is a fine choice for a representation scheme. However, typically in computing applications, it is more common to use the **floating point representation scheme** (see Fig. 2.5) to represent real numbers. In the floating point representation scheme we represent \( x \in \mathbb{R} \) by the pair \((b, e)\) of (positive or negative) integers of some prescribed sizes (determined by the desired accuracy) such that \( b \times 2^e \) is closest to \( x \). Floating point representation is the base-two version of **scientific notation**, where one represents a number \( y \in \mathbb{R} \) as its approximation of the form \( b \times 10^e \) for \( b, e \). It is called “floating
point” because we can think of the number \( b \) as specifying a sequence of binary digits, and \( e \) as describing the location of the “binary point” within this sequence. The use of floating representation is the reason why in many programming systems, printing the expression \( 0.1 + 0.2 \) will result in \( 0.30000000000000004 \) and not \( 0.3 \), see here, here and here for more.

The reader might be (rightly) worried about the fact that the floating point representation (or the rational number one) can only approximately represent real numbers. In many (though not all) computational applications, one can make the accuracy tight enough so that this does not affect the final result, though sometimes we do need to be careful. Indeed, floating-point bugs can sometimes be no joking matter. For example, floating point rounding errors have been implicated in the failure of a U.S. Patriot missile to intercept an Iraqi Scud missile, costing 28 lives, as well as a 100 million pound error in computing payouts to British pensioners.

2.3.1 Can we represent reals exactly?
Given the issues with floating point approximations for real numbers, a natural question is whether it is possible to represent real numbers exactly as strings. Unfortunately, the following theorem shows that this cannot be done:

**Theorem 2.5 — Reals are uncountable.** There does not exist a one-to-one function \( R \rightarrow S \) for real numbers to strings. \(^3\)

Theorem 2.5 was proven by Georg Cantor in 1874. (Cantor used the set \( \mathbb{N} \) rather than \( \{0, 1\}^* \), but one can show that these two results are equivalent using the one-to-one mappings between those two sets, see Exercise 2.13.) The non-existence of such a map is equivalent to saying that there is no way to “count” all the real numbers as some sequence \( x_0, x_1, x_2, \ldots \). For this reason Theorem 2.5 is known as the uncountability of the reals.

This result (and the theory around it) was quite shocking to mathematicians at the time. By showing that there is no one-to-one map from \( \mathbb{R} \) to \( \{0, 1\}^* \) (or \( \mathbb{N} \)), Cantor showed that these two infinite sets have “different forms of infinity” and that the set of real numbers \( \mathbb{R} \) is in some sense “bigger” than the infinite set \( \{0, 1\}^* \). The notion that there are “shades of infinity” was deeply disturbing to mathematicians and philosophers at the time. The philosopher Ludwig Wittgenstein (whom we mentioned before) called Cantor’s results “utter nonsense” and “laughable.” Others thought they were even worse than that. Leopold Kronecker called Cantor a “corrupter of youth,” while Henri Poincaré said that Cantor’s ideas “should be banished from mathematics once and for all.” The tide eventually turned,

\(^3\) \( R \rightarrow S \) stands for “real numbers to strings”. 

---

**Figure 2.6:** XKCD cartoon on floating-point arithmetic.
and these days Cantor’s work is universally accepted as the cornerstone of set theory and the foundations of mathematics. As David Hilbert said in 1925, “No one shall expel us from the paradise which Cantor has created for us.” As we will see later in this book, Cantor’s ideas also play a huge role in the theory of computation.

Now that we have discussed Theorem 2.5’s importance, let us see the proof. It is achieved in two steps:

1. Define some infinite set \( \mathcal{X} \) for which it is easier for us to prove that \( \mathcal{X} \) is not countable (namely, it’s easier for us to prove there is no one-to-one function from \( \mathcal{X} \) to \( \{0, 1\}^* \)).

2. Prove that there is a one-to-one function \( G \) mapping \( \mathcal{X} \) to \( \mathbb{R} \).

We can use a proof by contradiction to show that these two facts together imply Theorem 2.5. Specifically, if we assume (towards the sake of contradiction) that there exists some one-to-one \( F \) mapping \( \mathbb{R} \) to \( \{0, 1\}^* \) then the function \( x \mapsto F(G(x)) \) obtained by composing \( F \) with the function \( G \) from Step 2 above would be a one-to-one function from \( \mathcal{X} \) to \( \{0, 1\}^* \), which contradicts what we proved in Step 1!

To turn this idea into a full proof of Theorem 2.5 we need to:

- Define the set \( \mathcal{X} \).
- Prove that there is no one-to-one function from \( \mathcal{X} \) to \( \{0, 1\}^* \).
- Prove that there is a one-to-one function from \( \mathcal{X} \) to \( \mathbb{R} \).

We now proceed to do precisely that. That is, we will define the set \( \{0, 1\}^\infty \), which will play the role of \( \mathcal{X} \), and then state and prove two lemmas that show that this set satisfies our two desired properties.

**Definition 2.6** We denote by \( \{0, 1\}^\infty \) the set \( \{ f \mid f : \mathbb{N} \rightarrow \{0, 1\} \} \).

That is, \( \{0, 1\}^\infty \) is a set of functions, and a function \( f \) is in \( \{0, 1\}^\infty \) iff its domain is \( \mathbb{N} \) and its codomain is \( \{0, 1\} \). We can also think of \( \{0, 1\}^\infty \) as the set of all infinite sequences of bits, since a function \( f : \mathbb{N} \rightarrow \{0, 1\} \) can be identified with the sequence \((f(0), f(1), f(2), \ldots)\). The following two lemmas show that \( \{0, 1\}^\infty \) can play the role of \( \mathcal{X} \) to establish Theorem 2.5.

**Lemma 2.7** There does not exist a one-to-one map \( \text{FtS} : \{0, 1\}^\infty \rightarrow \{0, 1\}^* \).

**Lemma 2.8** There does exist a one-to-one map \( \text{FtR} : \{0, 1\}^\infty \rightarrow \mathbb{R} \).

As we’ve seen above, Lemma 2.7 and Lemma 2.8 together imply Theorem 2.5. To repeat the argument more formally, suppose, for the sake of contradiction, that there did exist a one-to-one function

\(^4\) \textit{FtS} stands for “functions to strings”.

\(^5\) \textit{FtR} stands for “functions to reals.”
$RtS : \mathbb{R} \to \{0, 1\}^*$. By Lemma 2.8, there exists a one-to-one function $FtR : \{0, 1\}^\infty \to \mathbb{R}$. Thus, under this assumption, since the composition of two one-to-one functions is one-to-one (see Exercise 2.12), the function $FtS : \{0, 1\}^\infty \to \{0, 1\}^*$ defined as $FtS(f) = RtS(FtR(f))$ will be one to one, contradicting Lemma 2.7. See Fig. 2.7 for a graphical illustration of this argument.

Now all that is left is to prove these two lemmas. We start by proving Lemma 2.7 which is really the heart of Theorem 2.5.

**Proof of Lemma 2.7.** We will prove that there does not exist an onto function $StF : \{0, 1\}^* \to \{0, 1\}^\infty$. This implies the lemma since for every two sets $A$ and $B$, there exists an onto function from $A$ to $B$.
The technique of this proof is known as the “diagonal argument” and is illustrated in Fig. 2.8. We assume, towards a contradiction, that there exists such a function $\text{St}F : \{0, 1\}^* \rightarrow \{0, 1\}^\infty$. We will show that $\text{St}F$ is not onto by demonstrating a function $\overline{d} \in \{0, 1\}^\infty$ such that $\overline{d} \neq \text{St}F(x)$ for every $x \in \{0, 1\}^*$. Consider the lexicographic ordering of binary strings (i.e., “”, 0, 1, 00, 01, …). For every $n \in \mathbb{N}$, we let $x_n$ be the $n$-th string in this order. That is $x_0 = “$, $x_1 = 0$, $x_2 = 1$ and so on and so forth. We define the function $\overline{d} \in \{0, 1\}^\infty$ as follows:

$$\overline{d}(n) = 1 - \text{St}F(x_n)(n) \quad (2.4)$$

for every $n \in \mathbb{N}$. That is, to compute $\overline{d}$ on input $n \in \mathbb{N}$, we first compute $g = \text{St}F(x_n)$, where $x_n \in \{0, 1\}^*$ is the $n$-th string in the lexicographical ordering. Since $g \in \{0, 1\}^\infty$, it is a function mapping $\mathbb{N}$ to $\{0, 1\}$. The value $\overline{d}(n)$ is defined to be the negation of $g(n)$.

The definition of the function $\overline{d}$ is a bit subtle. One way to think about it is to imagine the function $\text{St}F$ as being specified by an infinitely long table, in which every row corresponds to a string $x \in \{0, 1\}^*$ (with strings sorted in lexicographic order), and contains the sequence $\text{St}F(x)(0), \text{St}F(x)(1), \text{St}F(x)(2), …$. The diagonal elements in this table are the values

$$\text{St}F(“”)(0), \text{St}F(0)(1), \text{St}F(1)(2), \text{St}F(00)(3), \text{St}F(01)(4), … \quad (2.5)$$

which correspond to the elements $\text{St}F(x_n)(n)$ in the $n$-th row and $n$-th column of this table for $n = 0, 1, 2, …$. The function $\overline{d}$ we defined above maps every $n \in \mathbb{N}$ to the negation of the $n$-th diagonal value.

To complete the proof that $\text{St}F$ is not onto we need to show that $\overline{d} \neq \text{St}F(x)$ for every $x \in \{0, 1\}^*$. Indeed, let $x \in \{0, 1\}^*$ be some string and let $g = \text{St}F(x)$. If $n$ is the position of $x$ in the lexicographical order then by construction $\overline{d}(n) = 11 - g(n) \neq g(n)$ which means that $g \neq \overline{d}$ which is what we wanted to prove.

The proof of Lemma 2.7 is rather subtle, and worth re-reading a second or third time. We will use the “diagonal argument” again several times later on in this book.
Remark 2.9 — Generalizing beyond strings and reals.

Lemma 2.7 doesn’t really have much to do with the natural numbers or the strings. An examination of the proof shows that it really shows that for every set \( S \), there is no one-to-one map \( F : \{0, 1\}^S \to S \) where \( \{0, 1\}^S \) denotes the set \( \{ f \mid f : S \to \{0, 1\} \} \) of all Boolean functions with domain \( S \). Since we can identify a subset \( V \subseteq S \) with its characteristic function \( f = 1_V \) (i.e., \( 1_V(x) = 1 \) iff \( x \in V \)), we can think of \( \{0, 1\}^S \) also as the set of all subsets of \( S \). This subset is sometimes called the power set of \( S \).

The proof of Lemma 2.7 can be generalized to show that there is no one-to-one map between a set and its power set. In particular, it means that the set \( \{0, 1\}^\mathbb{R} \) is “even bigger” than \( \mathbb{R} \). Cantor used these ideas to construct an infinite hierarchy of shades of infinity. The number of such shades turns out to be much larger than \( |\mathbb{N}| \) or even \( |\mathbb{R}| \). He denoted the cardinality of \( \mathbb{N} \) by \( \aleph_0 \) and denoted the next largest infinite number by \( \aleph_1 \). (\( \aleph \) is the first letter in the Hebrew alphabet.) Cantor also made the continuum hypothesis that \( |\mathbb{R}| = \aleph_1 \). We will come back to the fascinating story of this hypothesis later on in this book. This lecture of Aaronson mentions some of these issues (see also this Berkeley CS 70 lecture).

To complete the proof of Theorem 2.5, we need to show Lemma 2.8. This requires some calculus background but is otherwise straightforward. If you have not had much experience with limits of a real series before, then the formal proof below might be a little hard to follow. This part is not the core of Cantor’s argument, nor are such limits important to the remainder of this book, so you can feel free to take Lemma 2.8 on faith and skip the proof.

Proof Idea:

We define \( FtR(f) \) to be the number between 0 and 2 whose decimal expansion is \( f(0).f(1)f(2) \ldots \), or in other words \( FtR(f) = \sum_{i=0}^{\infty} f(i) \cdot 10^{-i} \). To prove that \( FtR \) is one to one, we need to show that if \( f \not= g \) then \( FtR(f) \not= FtR(g) \). To do that we let \( k \in \mathbb{N} \) be the first input on which \( f \) and \( g \) disagree. The numbers \( FtR(f) \) and \( FtR(g) \) agree in the first \( k - 2 \) digits following the decimal point and disagree in the \( k - 1 \)-th digit. One can then calculate and verify that this means that \(|FtR(f) - FtR(g)| > 0.5 \cdot 10^{-k} \) which in particular means that these two numbers are distinct from one another. (You might wonder why we can’t immediately deduce that two numbers that differ in a digit are not the same. The issue is that we have to be a little more careful when talking about infinite expansions. For example, the number one half has two decimal expansions 0.5 and 0.49999⋯. However, this
issue does not come up if (as in our case) we restrict attention only to numbers with decimal expansions that do not involve the digit 9.

* 

Proof of Lemma 2.8. For every \( f \in \{0, 1\}^\infty \), we define \( FtR(f) \) to be the number whose decimal expansion is \( f(0).f(1)f(2)f(3) \ldots \).

Formally we define
\[
FtR(f) = \sum_{i=0}^{\infty} f(i) \cdot 10^{-i} \tag{2.6}
\]

It is a known result in calculus (whose proof we will not repeat here) that the series on the righthand side of (2.7) converges to a definite limit in \( \mathbb{R} \).

We now prove that \( FtR \) is one to one. Let \( f, g \) be two distinct functions in \( \{0, 1\}^\infty \). Since \( f \) and \( g \) are distinct, there must be some input on which they differ, and we define \( k \) to be the smallest such input and assume without loss of generality that \( f(k) = 0 \) and \( g(k) = 1 \).

(Otherwise, if \( f(k) = 1 \) and \( g(k) = 0 \), then we can simply switch the roles of \( f \) and \( g \).) The numbers \( FtR(f) \) and \( FtR(g) \) agree with each other up to the \( k-1 \)-th digit up after the decimal point. Since this digit equals 0 for \( FtR(f) \) and equals 1 for \( FtR(g) \), we claim that \( FtR(g) \) is bigger than \( FtR(f) \) by at least \( 0.5 \cdot 10^{-k} \). To see this note that the difference \( FtR(g) - FtR(f) \) will be minimized if \( g(\ell) = 0 \) for every \( \ell > k \) and \( f(\ell) = 1 \) for every \( \ell > k \), in which case (since \( f \) and \( g \) agree up to the \( k-1 \)-th digit)
\[
FtR(g) - FtR(f) = 10^{-k} - 10^{-k-1} - 10^{-k-2} - 10^{-k-3} - \cdots \tag{2.7}
\]

Since the infinite series \( \sum_{j=0}^{\infty} 10^{-i} \) converges to \( 10/9 \), it follows that for every such \( f \) and \( g \), \( FtR(g) - FtR(f) \geq 10^{-k} - 10^{-k-1} \cdot (10/9) > 0 \). In particular we see that for every distinct \( f, g \in \{0, 1\}^\infty \), \( FtR(f) \neq FtR(g) \), implying that the function \( FtR \) is one to one.

\[
\boxed{
\text{Remark 2.10 — Using decimal expansion (optional). In the proof above we used the fact that}
\text{1 + 1/10 + 1/100 + \cdots converges to 10/9, which plugging into (2.7) yields that the difference between}
\text{\(FtR(g)\) and \(FtR(h)\) is at least \(10^{-k} - 10^{-k-1} \cdot (10/9) > 0\). While the choice of the decimal representation for \(FtR\)
\text{was arbitrary, we could not have used the binary representation in its place. Had we used the binary expansion instead of decimal, the corresponding sequence 1 + 1/2 + 1/4 + \cdots converges to \(2/1 = 2\).}
}
since \(2^{-k} = 2^{-k-1} \cdot 2\), we could not have deduced that \(F/R\) is one to one. Indeed there do exist pairs of distinct sequences \(f, g \in \{0, 1\}^\infty\) such that \(\sum_{i=0}^\infty f(i)2^{-i} = \sum_{i=0}^\infty g(i)2^{-i}\). (For example, the sequence 1, 0, 0, 0, ... and the sequence 0, 1, 1, ... have this property.)

### 2.4 Representing Objects Beyond Numbers

Numbers are of course by no means the only objects that we can represented as binary strings. A representation scheme for representing objects from some set \(\mathcal{O}\) consists of an encoding function that maps an object in \(\mathcal{O}\) to a string, and a decoding function that decodes a string back to an object in \(\mathcal{O}\). Formally, we make the following definition:

**Definition 2.11 — String representation.** Let \(\mathcal{O}\) be any set. A representation scheme for \(\mathcal{O}\) is a pair of functions \(E, D\) where \(E : \mathcal{O} \to \{0, 1\}^*\) is a total one-to-one function, \(D : \{0, 1\}^* \to \mathcal{O}\) is a (possibly partial) function, and such that \(D(E(o)) = o\) for every \(o \in \mathcal{O}\). \(E\) is known as the encoding function and \(D\) is known as the decoding function.

Note that the condition \(D(E(o)) = o\) for every \(o \in \mathcal{O}\) implies that \(D\) is onto (can you see why?). It turns out that to construct a representation scheme we only need to find an encoding function. That is, every one-to-one encoding function has a corresponding decoding function, as shown in the following lemma:

**Lemma 2.12** Suppose that \(E : \mathcal{O} \to \{0, 1\}^*\) is one-to-one. Then there exists a function \(D : \{0, 1\}^* \to \mathcal{O}\) such that \(D(E(o)) = o\) for every \(o \in \mathcal{O}\).

**Proof.** Let \(o_0\) be some arbitrary element of \(\mathcal{O}\). For every \(x \in \{0, 1\}^*\), there exists either zero or a single \(o \in \mathcal{O}\) such that \(E(o) = x\) (otherwise \(E\) would not be one-to-one). We will define \(D(x)\) to equal \(o_0\) in the first case and this single object \(o\) in the second case. By definition \(D(E(o)) = o\) for every \(o \in \mathcal{O}\).

**Remark 2.13 — Total decoding functions.** While the decoding function of a representation scheme can in general be a partial function, the proof of Lemma 2.12 implies that every representation scheme has a total decoding function. This observation can sometimes be useful.
2.4.1 Finite representations

If \( \mathcal{O} \) is finite, then we can represent every object in \( \mathcal{O} \) as a string of length at most some number \( n \). What is the value of \( n \)? Let us denote by \( \{0, 1\}^n \) the set \( \{ x \in \{0, 1\}^* : |x| \leq n \} \) of strings of length at most \( n \). The size of \( \{0, 1\}^n \) is equal to:

\[
|\{0, 1\}^0| + |\{0, 1\}^1| + |\{0, 1\}^2| + \cdots + |\{0, 1\}^n| = \sum_{i=0}^{n} 2^i = 2^{n+1} - 1 \quad (2.8)
\]

using the standard formula for summing a geometric progression.

To obtain a representation of objects in \( \mathcal{O} \) as strings of length at most \( n \) we need to come up with a one-to-one function from \( \mathcal{O} \) to \( \{0, 1\}^n \). We can do so, if and only if \(|\mathcal{O}| \leq 2^{n+1} - 1\) as is implied by the following lemma:

**Lemma 2.14** For every two finite sets \( S, T \), there exists a one-to-one function \( E : S \rightarrow T \) if and only if \(|S| \leq |T|\).

**Proof.** Let \( k = |S| \) and \( m = |T| \) and so write the elements of \( S \) and \( T \) as \( S = \{ s_0, s_1, \ldots, s_{k-1} \} \) and \( T = \{ t_0, t_1, \ldots, t_{m-1} \} \). We need to show that there is a one-to-one function \( E : S \rightarrow T \) iff \( k \leq m \). For the “if” direction, if \( k \leq m \) we can simply define \( E(s_i) = t_i \) for every \( i \in [k] \). Clearly for \( i \neq j, t_i = E(s_i) \neq E(s_j) = t_j \), and hence this function is one-to-one. In the other direction, suppose that \( k > m \) and \( E : S \rightarrow T \) is some function. Then \( E \) cannot be one-to-one. Indeed, for \( i = 0, 1, \ldots, m - 1 \) let us “mark” the element \( t_j = E(s_i) \) in \( T \). If \( t_j \) was marked before, then we have found two objects in \( S \) mapping to the same element \( t_j \). Otherwise, since \( T \) has \( m \) elements, when we get to \( i = m - 1 \) we mark all the objects in \( T \). Hence, in this case, \( E(s_m) \) must map to an element that was already marked before.\(^6\)

\[\square\]

2.4.2 Prefix-free encoding

When showing a representation scheme for rational numbers, we used the “hack” of encoding the alphabet \( \{0, 1, \|\} \) to represent tuples of strings as a single string. This is a special case of the general paradigm of prefix-free encoding. The idea is the following: if our representation has the property that no string \( x \) representing an object \( o \) is a prefix (i.e., an initial substring) of a string \( y \) representing a different object \( o' \), then we can represent a lists of objects by merely concatenating the representations of all the list members. For example, because in English every sentence ends with a punctuation mark such as a period, exclamation, or question mark, no sentence can be a prefix of another and so we can represent a list of sentences by merely concatenating the sentences one after the other.\(^7\)

It turns out that we can transform every representation to a prefix-free form. This justifies Big Idea 1, and allows us to transform a repre-

\(^6\) This observation is sometimes known as the “Pigeon Hole Principle”: the principle that if you have a pigeon coop with \( m \) holes, and \( k > m \) pigeons, then there must be two pigeons in the same hole.

\(^7\) English has some complications such as periods used for abbreviations (e.g., “e.g.”) or sentence quotes containing punctuation, but high level point of a prefix-free representation for sentences still holds.
representation scheme for objects of a type $T$ to a representation scheme of lists of objects of the type $T$. By repeating the same technique, we can also represent lists of lists of objects of type $T$, and so on and so forth. But first let us formally define prefix-freeness:

**Definition 2.15 — Prefix free encoding.** For two strings $y, y'$, we say that $y$ is a prefix of $y'$ if $|y| \leq |y'|$ and for every $i < |y|, y_i = y'_i$.

Let $\mathcal{O}$ be a non-empty set and $E : \mathcal{O} \rightarrow \{0,1\}^*$ be a function. We say that $E$ is prefix-free if $E(o)$ is non-empty for every $o \in \mathcal{O}$ and there does not exist a distinct pair of objects $o, o' \in \mathcal{O}$ such that $E(o)$ is a prefix of $E(o')$.

Recall that for every set $\mathcal{O}$, the set $\mathcal{O}^*$ consists of all finite length tuples (i.e., lists) of elements in $\mathcal{O}$. The following theorem shows that if $E$ is a prefix-free encoding of $\mathcal{O}$ then by concatenating encodings we can obtain a valid (i.e., one-to-one) representation of $\mathcal{O}^*$:

**Theorem 2.16 — Prefix-free implies tuple encoding.** Suppose that $E : \mathcal{O} \rightarrow \{0,1\}^*$ is prefix-free. Then the following map $\overline{E} : \mathcal{O}^{*} \rightarrow \{0,1\}^*$ is one to one, for every $o_0, \ldots, o_{k-1} \in \mathcal{O}^*$, we define

$$\overline{E}(o_0, \ldots, o_{k-1}) = E(o_0)E(o_1) \cdots E(o_{k-1}).$$

(2.9)

**Proof Idea:**

The idea behind the proof is simple. Suppose that for example we want to decode a triple $(o_0, o_1, o_2)$ from its representation $x = \overline{E}(o_0, o_1, o_2) = E(o_0)E(o_1)E(o_2)$. We will do so by first finding the first prefix $x_0$ of $x$ such is a representation of some object. Then we will decode this object, remove $x_0$ from $x$ to obtain a new string $x'$, and continue onwards to find the first prefix $x_1$ of $x'$ and so on and so forth (see Exercise 2.9). The prefix-freeness property of $E$ will ensure that $x_0$ will in fact be $E(o_0), x_1$ will be $E(o_1)$, etc.

*
**Proof of Theorem 2.16.** We now show the formal proof. Suppose, towards the sake of contradiction, that there exist two distinct tuples \((o_0, \ldots, o_{k-1})\) and \((o'_0, \ldots, o'_{k'-1})\) such that
\[
E(o_0, \ldots, o_{k-1}) = E(o'_0, \ldots, o'_{k'-1}).
\] (2.10)

We will denote the string \(E(o_0, \ldots, o_{k-1})\) by \(x\).

Let \(i\) be the first index such that \(o_i \neq o'_i\). (If \(o_i = o'_i\) for all \(i\) then, since we assume the two tuples are distinct, one of them must be larger than the other. In this case we assume without loss of generality that \(k' > k\) and let \(i = k\).) In the case that \(i < k\), we see that the string \(x\) can be written in two different ways:
\[
x = E(o_0, \ldots, o_{k-1}) = x_0 \cdots x_i E(o_i) E(o_{i+1}) \cdots E(o_{k-1})
\] (2.11)
and
\[
x = E(o'_0, \ldots, o'_{k'-1}) = x_0 \cdots x_i E(o'_i) E(o'_{i+1}) \cdots E(o'_{k'-1})
\] (2.12)

where \(x_j = E(o_j) = E(o'_j)\) for all \(j < i\). Let \(\overline{y}\) be the string obtained after removing the prefix \(x_0 \cdots x_i\) from \(\overline{x}\). We see that \(\overline{y}\) can be written as both \(\overline{y} = E(o_j)s\) for some string \(s \in \{0, 1\}^*\) and as \(\overline{y} = E(o'_j)s'\) for some \(s' \in \{0, 1\}^*\). But this means that one of \(E(o_i)\) and \(E(o'_i)\) must be a prefix of the other, contradicting the prefix-freeness of \(E\).

In the case that \(i = k\) and \(k' > k\), we get a contradiction in the following way. In this case
\[
x = E(o_0) \cdots E(o_{k-1}) = E(o_0) \cdots E(o_{k-1}) E(o'_k) \cdots E(o'_{k'-1})
\] (2.13)

which means that \(E(o'_k) \cdots E(o'_{k'-1})\) must correspond to the empty string "". But in such a case \(E(o'_k)\) must be the empty string, which in particular is the prefix of any other string, contradicting the prefix-freeness of \(E\).

\[\square\]

**Remark 2.17 — Prefix freeness of list representation.**

Even if the representation \(E\) of objects in \(\mathcal{O}\) is prefix free, this does not mean that our representation \(\overline{E}\) of lists of such objects will be prefix free as well. In fact, it won’t be: for every three objects \(o, o', o''\) the representation of the list \((o, o')\) will be a prefix of the representation of the list \((o, o', o'')\). However, as we see in Lemma 2.18 below, we can transform every representation into prefix-free form, and so will be able to
2.4.3 Making representations prefix-free

Some natural representations are prefix-free. For example, every fixed output length representation (i.e., one-to-one function $E : \mathcal{O} \rightarrow \{0, 1\}^n$) is automatically prefix-free, since a string $x$ can only be a prefix of an equal-length $x'$ if $x$ and $x'$ are identical. Moreover, the approach we used for representing rational numbers can be used to show the following:

**Lemma 2.18** Let $E : \mathcal{O} \rightarrow \{0, 1\}^*$ be a one-to-one function. Then there is a one-to-one prefix-free encoding $\overline{E}$ such that $|\overline{E}(o)| \leq 2|E(o)| + 2$ for every $o \in \mathcal{O}$.

**Proof of Lemma 2.18.** The idea behind the proof is to use the map $0 \mapsto 00, 1 \mapsto 11$ to “double” every bit in the string $x$ and then mark the end of the string by concatenating to it the pair $01$. If we encode a string $x$ in this way, it ensures that the encoding of $x$ is never a prefix of the encoding of a distinct string $x'$. Formally, we define the function $PF : \{0, 1\}^* \rightarrow \{0, 1\}^*$ as follows:

$$PF(x) = x_0x_1x_2 \ldots x_{n-1}x_{n-1}01 \quad (2.14)$$

for every $x \in \{0, 1\}^*$. If $E : \mathcal{O} \rightarrow \{0, 1\}^*$ is the (potentially not prefix-free) representation for $\mathcal{O}$, then we transform it into a prefix-free representation $\overline{E} : \mathcal{O} \rightarrow \{0, 1\}^*$ by defining $\overline{E}(o) = PF(E(o))$.

To prove the lemma we need to show that (1) $\overline{E}$ is one-to-one and (2) $\overline{E}$ is prefix-free. In fact, prefix freeness is a stronger condition than one-to-one (if two strings are equal then in particular one of them is a prefix of the other) and hence it suffices to prove (2), which we now do.

Let $o \neq o'$ in $\mathcal{O}$ be two distinct objects. We will prove that $\overline{E}(o)$ is not a prefix of $\overline{E}(o')$. Define $x = E(o)$ and $x' = E(o')$. Since $E$ is one-to-one, $x \neq x'$. Under our assumption, $PF(x)$ is a prefix of $PF(x')$. If $|x| < |x'|$ then the two bits in positions $2|x|, 2|x| + 1$ in $PF(x)$ have the value $01$ but the corresponding bits in $PF(x')$ will equal either $00$ or $11$ (depending on the $|x|$-th bit of $x'$) and hence
PF(x) cannot be a prefix of PF(x'). If |x| = |x'| then, since x ≠ x', there must be a coordinate i in which they differ, meaning that the strings PF(x) and PF(x') differ in the coordinates 2i, 2i + 1, which again means that PF(x) cannot be a prefix of PF(x'). If |x| > |x'| then |PF(x)| = 2|x| + 2 > |PF(x')| = 2|x'| + 2 and hence PF(x) is longer than (and cannot be a prefix of) PF(x'). In all cases we see that PF(x) = \overline{E}(o) is not a prefix of PF(x') = \overline{E}(o'), hence completing the proof.

The proof of Lemma 2.18 is not the only or even the best way to transform an arbitrary representation into prefix-free form. Exercise 2.10 asks you to construct a more efficient prefix-free transformation satisfying |E(o)| ≤ |E(o)| + O(log |E(o)|).

2.4.4 “Proof by Python” (optional)
The proofs of Theorem 2.16 and Lemma 2.18 are constructive in the sense that they give us:

- A way to transform the encoding and decoding functions of any representation of an object O to an encoding and decoding functions that are prefix-free, and

- A way to extend prefix-free encoding and decoding of single objects to encoding and decoding of lists of objects by concatenation.

Specifically, we could transform any pair of Python functions encode and decode to functions pfencode and pfdecode that correspond to a prefix-free encoding and decoding. Similarly, given pfencode and pfdecode for single objects, we can extend them to encoding of lists. Let us show how this works for the case of the NtS and StN functions we defined above.

We start with the “Python proof” of Lemma 2.18: a way to transform an arbitrary representation into one that is prefix free. The function prefixfree below takes as input a pair of encoding and decoding functions, and returns a triple of functions containing prefix-free encoding and decoding functions, as well as a function that checks whether a string is a valid encoding of an object.

```python
# takes functions encode and decode mapping # objects to lists of bits and vice versa, # and returns functions pfencode and pfdecode that # maps objects to lists of bits and vice versa # in a prefix-free way. # Also returns a function pfvalid that says # whether a list is a valid encoding
```
def prefixfree(encode, decode):
    def pfencode(o):
        L = encode(o)
        return [L[i//2] for i in range(2*len(L))]+[0,1]
    def pfdecode(L):
        return decode([L[j] for j in range(0,len(L)-2,2)])
    def pfvalid(L):
        return (len(L) % 2 == 0 ) and all(L[2*i]==L[2*i+1] for i in range((len(L)-2)//2)) and L[-2:]==[0,1]
    return pfencode, pfdecode, pfvalid

pfNtS, pfStN, pfvalidN = prefixfree(NtS,StN)

NtS(234)
# 11101010
pfNtS(234)
# 111111001100110001
pfStN(pfNtS(234))
# 234
pfvalidM(pfNtS(234))
# true

We now show a “Python proof” of Theorem 2.16. Namely, we show a function represlists that takes as input a prefix-free representation scheme (implemented via encoding, decoding, and validity testing functions) and outputs a representation scheme for lists of such objects. If we want to make this representation prefix-free then we could fit it into the function prefixfree above.

def represlists(pfencode,pfdecode,pfvalid):
    """
    Takes functions pfencode, pfdecode and pfvalid, and returns functions encodelists, decodelists that can encode and decode lists of the objects
    """

8 When it's not too awkward, we use the term “Python function” or “subroutine” to distinguish between such snippets of Python programs and mathematical functions. However, in comments in Python source we use “functions” to denote Python functions, just as we use “integers” to denote Python int objects.
respectively.

""

def encodelist(L):
    """Gets list of objects, encodes it as list of bits""
    return ''.join([pfencode(obj) for obj in L])

def decodelist(S):
    """Gets lists of bits, returns lists of objects""
    i=0; j=1 ; res = []
    while j<=len(S):
        if pfvalid(S[i:j]):
            res += [pfdecode(S[i:j])]
            i=j
        j+= 1
    return res

return encodelist,decodelist

LtS, StL = represlists(pfNtS,pfStN,pfvalidN)

LtS([234,12,5])
# 11111001100110001111000111001101
StL(LtS([234,12,5])))
# [234, 12, 5]

2.4.5 Representing letters and text

We can represent a letter or symbol by a string, and then if this representation is prefix-free, we can represent a sequence of symbols by merely concatenating the representation of each symbol. One such representation is the ASCII that represents 128 letters and symbols as strings of 7 bits. Since the ASCII representation is fixed-length, it is automatically prefix-free (can you see why?). Unicode is representation of (at the time of this writing) about 128,000 symbols as numbers (known as code points) between 0 and 1, 114, 111. There are several types of prefix-free representations of the code points, a popular one being UTF-8 that encodes every codepoint into a string of length between 8 and 32.

Example 2.19 — The Braille representation. The Braille system is another way to encode letters and other symbols as binary strings. Specifi-
ally, in Braille, every letter is encoded as a string in \( \{0, 1\}^6 \), which is written using indented dots arranged in two columns and three rows, see Fig. 2.10. (Some symbols require more than one six-bit string to encode, and so Braille uses a more general prefix-free encoding.)

The Braille system was invented in 1821 by Louis Braille when he was just 12 years old (though he continued working on it and improving it throughout his life). Braille was a French boy that lost his eyesight at the age of 5 as the result of an accident.

**Example 2.20 — Representing objects in C (optional).** We can use programming languages to probe how our computing environment represents various values. This is easiest to do in “unsafe” programming languages such as C that allow direct access to the memory.

Using a simple C program we have produced the following representations of various values. One can see that for integers, multiplying by 2 corresponds to a “left shift” inside each byte. In contrast, for floating point numbers, multiplying by two corresponds to adding one to the exponent part of the representation. In the architecture we used, a negative number is represented using the two’s complement approach. C represents strings in a prefix-free form by ensuring that a zero byte is at their end.

```c
int  2 : 00000010 00000000 00000000 00000000
int  4 : 00000100 00000000 00000000 00000000
int  513 : 00000001 00000010 00000000 00000000
long 513 : 00000001 00000010 00000000 00000000
  0000000 00000000 0000000 00000000
int -1 : 11111111 11111111 11111111 11111111
int -2 : 11111110 11111111 11111111 11111111
string Hello: 01001000 01100101 01101100 01101100
  01101111 00000000
string abcd : 01100001 01100101 01101100 01101100
  01101111 00000000
float 33.0 : 00000000 00000000 00000100 01000010
float 66.0 : 00000000 00000000 10000100 01000010
float 132.0: 00000000 00000000 00000100 01000011
double 132.0: 00000000 00000000 00000000 00000000
  0000000 1000000 0110000 0100000
```
2.4.6 Representing vectors, matrices, images

Once we can represent numbers and lists of numbers, then we can also represent vectors (which are just lists of numbers). Similarly, we can represent lists of lists, and thus, in particular, can represent matrices. To represent an image, we can represent the color at each pixel by a list of three numbers corresponding to the intensity of Red, Green and Blue. Thus an image of \( n \) pixels would be represented by a list of \( n \) such length-three lists. A video can be represented as a list of images.

2.4.7 Representing graphs

A graph on \( n \) vertices can be represented as an \( n \times n \) adjacency matrix whose \((i, j)\)th entry is equal to 1 if the edge \((i, j)\) is present and is equal to 0 otherwise. That is, we can represent an \( n \) vertex directed graph \( G = (V, E) \) as a string \( A \in \{0, 1\}^n \) such that \( A_{i,j} = 1 \) iff the edge \( i \to j \in E \). We can transform an undirected graph to a directed graph by replacing every edge \( \{i, j\} \) with both edges \( i \to j \) and \( j \to i \).

Another representation for graphs is the adjacency list representation. That is, we identify the vertex set \( V \) of a graph with the set \([n]\) where \( n = |V| \), and represent the graph \( G = (V, E) \) as a list of \( n \) lists, where the \( i \)-th list consists of the out-neighbors of vertex \( i \). The difference between these representations can be significant for some applications, though for us would typically be immaterial.

2.4.8 Representing lists and nested lists

If we have a way of representing objects from a set \( \mathcal{O} \) as binary strings, then we can represent lists of these objects by applying a prefix-free transformation. Moreover, we can use a trick similar to the above to handle nested lists. The idea is that if we have some representation \( E : \mathcal{O} \to \{0, 1\}^* \), then we can represent nested lists of items from \( \mathcal{O} \) using strings over the five element alphabet \( \Sigma = \{ 0, 1, [\;], \; \} \). For example, if \( o_1 \) is represented by \( \emptyset 0011 \), \( o_2 \) is represented by \( 10011 \), and \( o_3 \) is represented by \( 00111 \), then we can represent the nested list \((o_1, (o_2, o_3))\) as the string "[0011, [10011, 00111]]" over the alphabet \( \Sigma \). By encoding every element of \( \Sigma \) itself as a three-bit string, we can transform any representation for objects \( \mathcal{O} \) into a representation that enables representing (potentially nested) lists of these objects.

2.4.9 Notation

We will typically identify an object with its representation as a string. For example, if \( F : \{0, 1\}^* \to \{0, 1\}^* \) is some function that maps strings to strings and \( n \) is an integer, we might make statements such as "\( F(n) + 1 \) is prime" to mean that if we represent \( n \) as a string \( x \), then the integer \( m \) represented by the string \( F(x) \) satisfies that \( m + 1 \) is prime. (You can see how this convention of identifying objects with
their representation can save us a lot of cumbersome formalism.) Similarly, if \( x, y \) are some objects and \( F \) is a function that takes strings as inputs, then by \( F(x, y) \) we will mean the result of applying \( F \) to the representation of the ordered pair \((x, y)\). We use the same notation to invoke functions on \( k \)-tuples of objects for every \( k \).

This convention of identifying an object with its representation as a string is one that we humans follow all the time. For example, when people say a statement such as “17 is a prime number”, what they really mean is that the integer whose decimal representation is the string “17”, is prime.

When we say

\[ A \text{ is an algorithm that computes the multiplication function on natural numbers.} \]

what we really mean is that

\[ A \text{ is an algorithm that computes the function } \]

\[ F : \{0, 1\}^* \rightarrow \{0, 1\}^* \text{ such that for every pair } a, b \in \mathbb{N}, \]

if \( x \in \{0, 1\}^* \) is a string representing the pair \((a, b)\) then \( F(x) \) will be a string representing their product \( a \cdot b \).

### 2.5 Defining Computational Tasks as Mathematical Functions

Abstractly, a computational process is some process that takes an input which is a string of bits and produces an output which is a string of bits. This transformation of input to output can be done using a modern computer, a person following instructions, the evolution of some natural system, or any other means.

In future chapters, we will turn to mathematically defining computational processes, but, as we discussed above, at the moment we focus on computational tasks. That is, we focus on the specification and not the implementation. Again, at an abstract level, a computational task can specify any relation that the output needs to have with the input. However, for most of this book, we will focus on the simplest and most common task of computing a function. Here are some examples:

- Given (a representation of) two integers \( x, y \), compute the product \( x \times y \). Using our representation above, this corresponds to computing a function from \( \{0, 1\}^* \) to \( \{0, 1\}^* \). We have seen that there is more than one way to solve this computational task, and in fact, we still do not know the best algorithm for this problem.

- Given (a representation of) an integer \( z > 1 \), compute its factorization; i.e., the list of primes \( p_1 \leq \cdots \leq p_k \) such that \( z = p_1 \cdots p_k \). This
again corresponds to computing a function from $\{0, 1\}^*$ to $\{0, 1\}^*$. The gaps in our knowledge of the complexity of this problem are even larger.

- Given (a representation of) a graph $G$ and two vertices $s$ and $t$, compute the length of the shortest path in $G$ between $s$ and $t$, or do the same for the longest path (with no repeated vertices) between $s$ and $t$. Both these tasks correspond to computing a function from $\{0, 1\}^*$ to $\{0, 1\}^*$, though it turns out that there is a vast difference in their computational difficulty.

- Given the code of a Python program, determine whether there is an input that would force it into an infinite loop. This task corresponds to computing a partial function from $\{0, 1\}^*$ to $\{0, 1\}$ since not every string corresponds to a syntactically valid Python program. We will see that we do understand the computational status of this problem, but the answer is quite surprising.

- Given (a representation of) an image $I$, decide if $I$ is a photo of a cat or a dog. This corresponds to computing some (partial) function from $\{0, 1\}^*$ to $\{0, 1\}$.

\begin{remark}
\textbf{Remark 2.21 — Boolean functions and languages.} An important special case of computational tasks corresponds to computing Boolean functions, whose output is a single bit $\{0, 1\}$. Computing such functions corresponds to answering a YES/NO question, and hence this task is also known as a decision problem. Given any function $F : \{0, 1\}^* \to \{0, 1\}$ and $x \in \{0, 1\}^*$, the task of computing $F(x)$ corresponds to the task of deciding whether or not $x \in L$ where $L = \{x : F(x) = 1\}$ is known as the language that corresponds to the function $F$. (The language terminology is due to historical connections between the theory of computation and formal linguistics as developed by Noam Chomsky.) Hence many texts refer to such a computational task as deciding a language.
\end{remark}

For every particular function $F$, there can be several possible algorithms to compute $F$. We will be interested in questions such as:

- For a given function $F$, can it be the case that there is no algorithm to compute $F$?

- If there is an algorithm, what is the best one? Could it be that $F$ is “effectively uncomputable” in the sense that every algorithm for computing $F$ requires a prohibitively large amount of resources?

\begin{figure}[h]
\centering
\includegraphics[width=0.4\textwidth]{fig2_12.png}
\caption{A subset $L \subseteq \{0, 1\}^*$ can be identified with the function $F : \{0, 1\}^* \to \{0, 1\}$ such that $F(x) = 1$ if $x \in L$ and $F(x) = 0$ if $x \notin L$. Functions with a single bit of output are called Boolean functions, while subsets of strings are called languages. The above shows that the two are essentially the same object, and we can identify the task of deciding membership in $L$ (known as deciding a language in the literature) with the task of computing the function $F$.}
\end{figure}
• If we cannot answer this question, can we show equivalence between different functions $F$ and $F'$ in the sense that either they are both easy (i.e., have fast algorithms) or they are both hard?

• Can a function being hard to compute ever be a good thing? Can we use it for applications in areas such as cryptography?

In order to do that, we will need to mathematically define the notion of an algorithm, which is what we will do in Chapter 3.

2.5.1 Distinguish functions from programs!

You should always watch out for potential confusions between specifications and implementations or equivalently between mathematical functions and algorithms/programs. It does not help that programming languages (Python included) use the term “functions” to denote (parts of) programs. This confusion also stems from thousands of years of mathematical history, where people typically defined functions by means of a way to compute them.

For example, consider the multiplication function on natural numbers. This is the function $\text{MULT} : \mathbb{N} \times \mathbb{N} \rightarrow \mathbb{N}$ that maps a pair $(x, y)$ of natural numbers to the number $x \cdot y$. As we mentioned, it can be implemented in more than one way:

```python
def mult1(x, y):
    res = 0
    while y>0:
        res += x
        y -= 1
    return res

def mult2(x, y):
    a = str(x)  # represent x as string in decimal notation
    b = str(y)  # represent y as string in decimal notation
    res = 0
    for i in range(len(a)):
        for j in range(len(b)):
            res += int(a[len(a)-i])*int(b[len(b)-j])*(10**(i+j))
    return res
```

```python
print(mult1(12,7))  # 84
print(mult2(12,7))  # 84
```

Both `mult1` and `mult2` produce the same output given the same pair of natural number inputs. (Though `mult1` will take far longer to
do so when the numbers become large.) Hence, even though these are two different programs, they compute the same mathematical function. This distinction between a program or algorithm \( A \), and the function \( F \) that \( A \) computes will be absolutely crucial for us in this course (see also Fig. 2.13).

**Big Idea 2** A function is not the same as a program. A program computes a function.

Distinguishing functions from programs (or other ways for computing, including circuits and machines) is a crucial theme for this course. For this reason, this is often a running theme in questions that I (and many other instructors) assign in homeworks and exams (hint, hint).

**Remark 2.22** — Computation beyond functions (advanced, optional). Functions capture quite a lot of computational tasks, but one can consider more general settings as well. For starters, we can and will talk about partial functions, that are not defined on all inputs. When computing a partial function, we only need to worry about the inputs on which the function is defined. Another way to say it is that we can design an algorithm for a partial function \( F \) under the assumption that someone “promised” us that all inputs \( x \) would be such that \( F(x) \) is defined (as otherwise, we do not care about the result). Hence such tasks are also known as promise problems.

Another generalization is to consider relations that may have more than one possible admissible output. For example, consider the task of finding any solution for a given set of equations. A relation \( R \) maps a string \( x \in \{0, 1\}^* \) into a set of strings \( R(x) \). For example, \( x \) might describe a set of equations, in which case \( R(x) \) would correspond to the set of all solutions to \( x \). We can also identify a relation \( R \) with the set of pairs of strings \( (x, y) \) where \( y \in R(x) \). A computational process solves a relation if for every \( x \in \{0, 1\}^* \), it outputs some string \( y \in R(x) \).

Later in this book, we will consider even more general tasks, including interactive tasks, such as finding a good strategy in a game, tasks defined using probabilistic notions, and others. However, for much of this book, we will focus on the task of computing a function, and often even a Boolean function, that has only a single bit of output. It turns out that a great deal of the theory of computation can be studied in the context of this task, and the insights learned are applicable in the more general settings.
Lecture Recap

- We can represent objects we want to compute on using binary strings.
- A representation scheme for a set of objects \( \mathcal{O} \) is a one-to-one map from \( \mathcal{O} \) to \( \{0, 1\}^* \).
- We can use prefix-free encoding to “boost” a representation for a set \( \mathcal{O} \) into representations of lists of elements in \( \mathcal{O} \).
- A basic computational task is the task of computing a function \( F : \{0, 1\}^* \rightarrow \{0, 1\}^* \). This task encompasses not just arithmetical computations such as multiplication, factoring, etc. but a great many other tasks arising in areas as diverse as scientific computing, artificial intelligence, image processing, data mining and many many more.
- We will study the question of finding (or at least giving bounds on) what is the best algorithm for computing \( F \) for various interesting functions \( F \).

2.6 Exercises

Exercise 2.1 Which one of these objects can be represented by a binary string?

a. An integer \( x \)

b. An undirected graph \( G \).

c. A directed graph \( H \)

d. All of the above.

Exercise 2.2 — Binary representation.

a. Prove that the function \( N tS : \mathbb{N} \rightarrow \{0, 1\}^* \) of the binary representation defined in (2.1) satisfies that for every \( n \in \mathbb{N} \), if \( x = N tS(n) \) then \( |x| = \lfloor \log_2 n \rfloor \) and \( x_i = \lfloor x/2^{|x|-i} \rfloor \mod 2 \).

b. Prove that \( N tS \) is a one to one function by coming up with a function \( StN : \{0, 1\}^* \rightarrow \mathbb{N} \) such that \( StN(N tS(n)) = n \) for every \( n \in \mathbb{N} \).

Exercise 2.3 — More compact than ASCII representation. The ASCII encoding can be used to encode a string of \( n \) English letters as a \( 7n \) bit binary string, but in this exercise, we ask about finding a more compact representation for strings of English lowercase letters.
1. Prove that there exists a representation scheme \((E, D)\) for strings over the 26-letter alphabet \(\{a, b, c, \ldots, z\}\) as binary strings such that for every \(n > 0\) and length-\(n\) string \(x \in \{a, b, \ldots, z\}^n\), the representation \(E(x)\) is a binary string of length at most \(4.8n + 1000\). In other words, prove that for every \(n\), there exists a one-to-one function \(E : \{a, b, \ldots, z\}^n \to \{0, 1\}^{\lfloor 4.8n+1000 \rfloor}\).

2. Prove that there exists no representation scheme for strings over the alphabet \(\{a, b, \ldots, z\}\) as binary strings such that for every length-\(n\) string \(x \in \{a, b, \ldots, z\}^n\), the representation \(E(x)\) is a binary string of length \(\lfloor 4.6n + 1000 \rfloor\). In other words, prove that there exists some \(n > 0\) such that there is no one-to-one function \(E : \{a, b, \ldots, z\}^n \to \{0, 1\}^{\lfloor 4.6n+1000 \rfloor}\).

3. Python’s `bz2.compress` function is a mapping from strings to strings, which uses the lossless (and hence one to one) bzip2 algorithm for compression. After converting to lowercase, and truncating spaces and numbers, the text of Tolstoy’s “War and Peace” contains \(n = 2,517,262\). Yet, if we run `bz2.compress` on the string of the text of “War and Peace” we get a string of length \(k = 6,274,768\) bits, which is only \(2.49n\) (and in particular much smaller than \(4.6n\)). Explain why this does not contradict your answer to the previous question.

4. Interestingly, if we try to apply `bz2.compress` on a random string, we get much worse performance. In my experiments, I got a ratio of about 4.78 between the number of bits in the output and the number of characters in the input. However, one could imagine that one could do better and that there exists a company called “Pied Piper” with an algorithm that can losslessly compress a string of \(n\) random lowercase letters to fewer than \(4.6n\) bits. Show that this is not the case by proving that for every \(n > 100\) and one to one function \(Encode : \{a, \ldots, z\}^n \to \{0, 1\}^*\), if we let \(Z\) be the random variable \(|Encode(x)|\) (i.e., the length of \(Encode(x)\)) for \(x\) chosen uniformly at random from the set \(\{a, \ldots, z\}^n\), then the expected value of \(Z\) is at least \(4.6n\).

---

**Exercise 2.4 — Representing graphs: upper bound.** Show that there is a string representation of directed graphs with vertex set \([n]\) and degree at most 10 that uses at most \(1000n \log n\) bits. More formally, show the following. Suppose we define for every \(n \in \mathbb{N}\), the set \(G_n\) as the set containing all directed graphs (with no self loops) over the vertex set \([n]\) where every vertex has degree at most 10. Then, prove that for every sufficiently large \(n\), there exists a one-to-one function \(E : G_n \to \{0, 1\}^{1000n \log n}\).
**Exercise 2.5 — Representing graphs: lower bound.** 1. Define $S_n$ to be the set of one-to-one and onto functions mapping $[n]$ to $[n]$. Prove that there is a one-to-one mapping from $S_n$ to $G_{2n}$, where $G_{2n}$ is the set defined in Exercise 2.4 above.

2. In this question you will show that one cannot improve the representation of Exercise 2.4 to length $o(n \log n)$. Specifically, prove for every sufficiently large $n \in \mathbb{N}$ there is no one-to-one function $E : G_n \to \{0, 1\}^{\ceil{0.001n \log n} + 1000}$.

**Exercise 2.6 — Multiplying in different representation.** Recall that the grade-school algorithm for multiplying two numbers requires $O(n^2)$ operations. Suppose that instead of using decimal representation, we use one of the following representations $R(x)$ to represent a number $x$ between 0 and $10^n - 1$. For which one of these representations you can still multiply the numbers in $O(n^2)$ operations?

a. The standard binary representation: $B(x) = (x_0, \ldots, x_k)$ where $x = \sum_{i=0}^{k} x_i 2^i$ and $k$ is the largest number s.t. $x \geq 2^k$.

b. The reverse binary representation: $B(x) = (x_k, \ldots, x_0)$ where $x_i$ is defined as above for $i = 0, \ldots, k - 1$.

c. Binary coded decimal representation: $B(x) = (y_0, \ldots, y_{n-1})$ where $y_i \in \{0, 1\}^4$ represents the $i^{th}$ decimal digit of $x$ mapping 0 to 0000, 1 to 0001, 2 to 0010, etc. (i.e. 9 maps to 1001)

d. All of the above.

**Exercise 2.7** Suppose that $R : \mathbb{N} \to \{0, 1\}^*$ corresponds to representing a number $x$ as a string of $x$ 1’s, (e.g., $R(4) = 1111$, $R(7) = 1111111$, etc.). If $x, y$ are numbers between 0 and $10^n - 1$, can we still multiply $x$ and $y$ using $O(n^2)$ operations if we are given them in the representation $R(\cdot)$?

**Exercise 2.8** Recall that if $F$ is a one-to-one and onto function mapping elements of a finite set $U$ into a finite set $V$ then the sizes of $U$ and $V$ are the same. Let $B : \mathbb{N} \to \{0, 1\}^*$ be the function such that for every $x \in \mathbb{N}$, $B(x)$ is the binary representation of $x$.

1. Prove that $x < 2^k$ if and only if $|B(x)| \leq k$.

2. Use 1. to compute the size of the set $\{y \in \{0, 1\}^* : |y| \leq k\}$ where $|y|$ denotes the length of the string $y$. 
 Exercise 2.9 — Prefix-free encoding of tuples. Suppose that $F : \mathbb{N} \to \{0, 1\}^*$ is a one-to-one function that is prefix-free in the sense that there is no $a \neq b$ s.t. $F(a)$ is a prefix of $F(b)$.

a. Prove that $F_2 : \mathbb{N} \times \mathbb{N} \to \{0, 1\}^*$, defined as $F_2(a, b) = F(a)F(b)$ (i.e., the concatenation of $F(a)$ and $F(b)$) is a one-to-one function.

b. Prove that $F_* : \mathbb{N}^* \to \{0, 1\}^*$ defined as $F_*(a_1, \ldots, a_k) = F(a_1)\ldots F(a_k)$ is a one-to-one function, where $\mathbb{N}^*$ denotes the set of all finite-length lists of natural numbers.

 Exercise 2.10 — More efficient prefix-free transformation. Suppose that $F : O \to \{0, 1\}^*$ is some (not necessarily prefix-free) representation of the objects in the set $O$, and $G : \mathbb{N} \to \{0, 1\}^*$ is a prefix-free representation of the natural numbers. Define $F'(o) = G(|F(o)|)F(o)$ (i.e., the concatenation of the representation of the length $F(o)$ and $F(o)$).

a. Prove that $F'$ is a prefix-free representation of $O$.

b. Show that we can transform any representation to a prefix-free one by a modification that takes a $k$ bit string into a string of length at most $k + O(\log k)$.

c. Show that we can transform any representation to a prefix-free one by a modification that takes a $k$ bit string into a string of length at most $k + \log k + O(\log \log k)$.

 Exercise 2.11 — Kraft’s Inequality. Suppose that $S \subseteq \{0, 1\}^n$ is some finite prefix-free set.

a. For every $k \leq n$ and length-$k$ string $x \in S$, let $L(x) \subseteq \{0, 1\}^n$ denote all the length-$n$ strings whose first $k$ bits are $x_0, \ldots, x_{k-1}$. Prove that (1) $|L(x)| = 2^{n-|x|}$ and (2) If $x \neq x'$ then $L(x)$ is disjoint from $L(x')$.

b. Prove that $\sum_{x \in S} 2^{-|x|} \leq 1$.

c. Prove that there is no prefix-free encoding of strings with less than logarithmic overhead. That is, prove that there is no function $PF : \{0, 1\}^* \to \{0, 1\}^*$ s.t. $|PF(x)| \leq |x| + 0.9 \log |x|$ for every $x \in \{0, 1\}^*$ and such that the set $\{PF(x) : x \in \{0, 1\}^*\}$ is prefix-free. The factor 0.9 is arbitrary; all that matters is that it is less than 1.
Exercise 2.12 — Composition of one-to-one functions. Prove that for every two one-to-one functions $F : S \to T$ and $G : T \to U$, the function $H : S \to U$ defined as $H(x) = G(F(x))$ is one to one.

Exercise 2.13 — Natural numbers and strings. 1. We have shown that the natural numbers can be represented as strings. Prove that the other direction holds as well: that there is a one-to-one map $StN : \{0, 1\}^* \to \mathbb{N}$. ($StN$ stands for “strings to numbers.”)

2. Recall that Cantor proved that there is no one-to-one map $RtN : \mathbb{R} \to \mathbb{N}$. Show that Cantor’s result implies Theorem 2.5.

Exercise 2.14 — Map lists of integers to a number. Recall that for every set $S$, the set $S^*$ is defined as the set of all finite sequences of members of $S$ (i.e., $S^* = \{(x_0, \ldots, x_{n-1}) \mid n \in \mathbb{N}, \forall i \in [n] x_i \in S\}$). Prove that there is a one-one-map from $\mathbb{Z}^*$ to $\mathbb{N}$ where $\mathbb{Z}$ is the set of \{-, -3, -2, -1, 0, +1, +2, +3, \ldots\} of all integers.

2.7 BIBLIOGRAPHICAL NOTES

The study of representing data as strings, including issues such as compression and error corrections falls under the purview of information theory, as covered in the classic textbook of Cover and Thomas [CT06]. Representations are also studied in the field of data structures design, as covered in texts such as [Cor+09].

The question of whether to represent integers with most significant digit first or last is known as Big Endian vs. Little Endian representation. This terminology comes from Cohen’s [Coh81] entertaining and informative paper about the conflict between adherents of both schools which he compared to the warring tribes in Jonathan Swift’s “Gulliver’s Travels”. The two’s complement representation of signed integers was suggested in von Neumann’s classic report [Neu45] that detailed the design approaches for a stored-program computer, though similar representations have been used even earlier in abacus and other mechanical computation devices.

The idea that we should separate the definition or specification of a function from its implementation or computation might seem “obvious,” but it took quite a lot of time for mathematicians to arrive at this viewpoint. Historically, a function $F$ was identified by rules or formulas showing how to derive the output from the input. As we discuss in greater depth in Chapter 8, in the 1800s this somewhat informal notion of a function started “breaking at the seams,” and eventually
mathematicians arrived at the more rigorous definition of a function as an arbitrary assignment of input to outputs. While many functions may be described (or computed) by one or more formulas, today we do not consider that to be an essential property of functions, and also allow functions that do not correspond to any “nice” formula.

We have mentioned that all representations of the real numbers are inherently approximate. Thus an important endeavor is to understand what guarantees we can offer on the approximation quality of the output of an algorithm, as a function of the approximation quality of the inputs. This question is known as the question of determining the numerical stability of given equations. The Floating Points Guide website contains an extensive description of the floating point representation, as well the many ways in which it could subtly fail, see also the website 0.30000000000000004.com.

Dauben [Dau90] gives a biography of Cantor with emphasis on the development of his mathematical ideas. [Hal60] is a classic textbook on set theory, including also Cantor’s theorem. Cantor’s Theorem is also covered in many texts on discrete mathematics, including [LehmanLeightonMeyer; LewisZax19].

The adjacency matrix representation of graphs is not merely a convenient way to map a graph into a binary string, but it turns out that many natural notions and operations on matrices are useful for graphs as well. (For example, Google’s PageRank algorithm relies on this viewpoint.) The notes of Spielman’s course are an excellent source for this area, known as spectral graph theory. We will return to this view much later in this book when we talk about random walks.
FINITE COMPUTATION
3

Defining computation

“there is no reason why mental as well as bodily labor should not be economized by the aid of machinery”, Charles Babbage, 1852

“If, unwarned by my example, any man shall undertake and shall succeed in constructing an engine embodying in itself the whole of the executive department of mathematical analysis upon different principles or by simpler mechanical means, I have no fear of leaving my reputation in his charge, for he alone will be fully able to appreciate the nature of my efforts and the value of their results.”, Charles Babbage, 1864

“To understand a program you must become both the machine and the program.”, Alan Perlis, 1982

People have been computing for thousands of years, with aids that include not just pen and paper, but also abacus, slide rules, various mechanical devices, and modern electronic computers. A priori, the notion of computation seems to be tied to the particular mechanism that you use. You might think that the “best” algorithm for multiplying numbers will differ if you implement it in Python on a modern laptop than if you use pen and paper. However, as we saw in the introduction (Chapter 0), an algorithm that is asymptotically better would eventually beat a worse one regardless of the underlying technology. This gives us hope for a technology independent way of defining computation. This is what we do in this chapter. We will define the notion of computing an output from an input by applying a sequence of basic operations (see Fig. 3.3). Using this, we will be able to precisely define statements such as “function $f$ can be computed by model $X$” or “function $f$ can be computed by model $X$ using $s$ operations”.

Learning Objectives:
- See that computation can be precisely modeled.
- Learn the computational model of Boolean circuits / straight-line programs.
- Equivalence of circuits and straight-line programs.
- Equivalence of AND/OR/NOT and NAND.
- Examples of computing in the physical world.

Figure 3.1: Calculating wheels by Charles Babbage. Image taken from the Mark I ‘operating manual’

Figure 3.2: A 1944 Popular Mechanics article on the Harvard Mark I computer.
3.1 DEFINING COMPUTATION

The name “algorithm” is derived from the Latin transliteration of Muhammad ibn Musa al-Khwarizmi’s name. Al-Khwarizmi was a Persian scholar during the 9th century whose books introduced the western world to the decimal positional numeral system, as well as to the solutions of linear and quadratic equations (see Fig. 3.4). However, Al-Khwarizmi’s descriptions of algorithms were rather informal by today’s standards. Rather than use “variables” such as $x, y$, he used concrete numbers such as 10 and 39, and trusted the reader to be able to extrapolate from these examples, much as algorithms are still taught to children today.

Here is how Al-Khwarizmi described the algorithm for solving an equation of the form $x^2 + bx = c$:

How to solve an equation of the form “roots and squares are equal to numbers”: For instance “one square, and ten roots of the same, amount to thirty-nine dirhems” that is to say, what must be the square which, when increased by ten of its own root, amounts to thirty-nine? The solution is this: you halve the number of the roots, which in the present instance yields five. This you multiply by itself; the product is twenty-five. Add this to thirty-nine; the sum is sixty-four. Now take the root of this, which is eight, and subtract from it half the number of roots, which is five; the remainder is three. This is the root of the square which you sought for; the square itself is nine.

For the purposes of this book, we will need a much more precise way to describe algorithms. Fortunately (or is it unfortunately?), at least at the moment, computers lag far behind school-age children
in learning from examples. Hence in the 20th century people came up with exact formalisms for describing algorithms, namely programming languages. Here is al-Khwarizmi’s quadratic equation solving algorithm described in the Python programming language:

```python
from math import sqrt
#Python speak to enable use of the sqrt function to compute square roots.
def solve_eq(b,c):
    # return solution of x^2 + bx = c following Al Kwarizmi's instructions
    # Al Kwarizmi demonstrates this for the case b=10 and c=39
    val1 = b / 2.0  # "halve the number of the roots"
    val2 = val1 * val1  # "this you multiply by itself"
    val3 = val2 + c  # "Add this to thirty-nine"
    val4 = sqrt(val3)  # "take the root of this"
    val5 = val4 - val1  # "subtract from it half the number of roots"
    return val5  # "This is the root of the square which you sought for"

# Test: solve x^2 + 10*x = 39
print(solve_eq(10,39))
# 3.0
```

We can define algorithms informally as follows:

**Informal definition of an algorithm**: An algorithm is a set of instructions for how to compute an output from an input by following a sequence of “elementary steps”. An algorithm $A$ computes a function $F$ if for every input $x$, if we follow the instructions of $A$ on the input $x$, we obtain the output $F(x)$.

In this chapter we will make this informal definition precise using the model of Boolean Circuits. We will show that Boolean Circuits are equivalent in power to straight line programs that are written in “ultra simple” programming languages that do not even have loops. We will also see that the particular choice of elementary operations is immaterial and many different choices yield models with equivalent power (see Fig. 3.6). However, it will take us some time to get there.
We will start by discussing what are “elementary operations” and how we map a description of an algorithm into an actual physical process that produces an output from an input in the real world.

### 3.2 COMPUTING USING AND, OR, AND NOT.

An algorithm breaks down a complex calculation into a series of simpler steps. These steps can be executed in a variety of different ways, including:

- Writing down symbols on a piece of paper.
- Modifying the current flowing on electrical wires.
- Binding a protein to a strand of DNA.
- Responding to a stimulus by a member of a collection (e.g., a bee in a colony, a trader in a market).

To formally define algorithms, let us try to “err on the side of simplicity” and model our “basic steps” as truly minimal. For example, here are some very simple functions:

- **OR** : \( \{0, 1\}^2 \rightarrow \{0, 1\} \) defined as

  \[
  OR(a, b) = \begin{cases} 
  0 & a = b = 0 \\ 
  1 & \text{otherwise}
  \end{cases}
  \]  

  (3.1)

- **AND** : \( \{0, 1\}^2 \rightarrow \{0, 1\} \) defined as

  \[
  AND(a, b) = \begin{cases} 
  1 & a = b = 1 \\ 
  0 & \text{otherwise}
  \end{cases}
  \]  

  (3.2)
• NOT : \{0, 1\} \rightarrow \{0, 1\} defined as

\[
\text{NOT}(a) = \begin{cases} 
0 & a = 1 \\
1 & a = 0 
\end{cases} \quad (3.3)
\]

The functions AND, OR and NOT, are the basic logical operators used in logic and many computer systems. In the context of logic, it is common to use the notation \(a \land b\) for AND\((a, b)\), \(a \lor b\) for OR\((a, b)\) and \(\overline{a}\) and \(\neg a\) for NOT\((a)\), and we will use this notation as well.

Each one of the functions AND, OR, NOT takes either one or two single bits as input, and produces a single bit as output. Clearly, it cannot get much more basic than that. However, the power of computation comes from composing such simple building blocks together.

**Example 3.1 — Majority from AND, OR and NOT.** Consider the function MAJ : \{0, 1\}^3 \rightarrow \{0, 1\} that is defined as follows:

\[
\text{MAJ}(x) = \begin{cases} 
1 & x_0 + x_1 + x_2 \geq 2 \\
0 & \text{otherwise} 
\end{cases} \quad (3.4)
\]

That is, for every \(x \in \{0, 1\}^3\), MAJ\((x) = 1\) if and only if the majority (i.e., at least two out of the three) of \(x\)'s elements are equal to 1. Can you come up with a formula involving AND, OR and NOT to compute MAJ? (It would be useful for you to pause at this point and work out the formula for yourself. As a hint, although the NOT operator is needed to compute some functions, you will not need to use it to compute MAJ.)

Let us first try to rephrase MAJ\((x)\) in words: “MAJ\((x) = 1\) if and only if there exists some pair of distinct elements \(i, j\) such that both \(x_i\) and \(x_j\) are equal to 1.” In other words it means that MAJ\((x) = 1\) iff either both \(x_0 = 1\) and \(x_1 = 1\), or both \(x_1 = 1\) and \(x_2 = 1\), or both \(x_0 = 1\) and \(x_2 = 1\). Since the OR of three conditions \(c_0, c_1, c_2\) can be written as OR\((c_0, OR(c_1, c_2))\), we can now translate this into a formula as follows:

\[
\text{MAJ}(x_0, x_1, x_2) = \text{OR} \left( \text{AND}(x_0, x_1), \text{OR}(\text{AND}(x_1, x_2), \text{AND}(x_0, x_2)) \right) . \quad (3.5)
\]

Recall that we can also write \(a \lor b\) for OR\((a, b)\) and \(a \land b\) for AND\((a, b)\). With this notation, (3.5) can also be written as

\[
\text{MAJ}(x_0, x_1, x_2) = ((x_0 \land x_1) \lor (x_1 \land x_2)) \lor (x_0 \land x_2) . \quad (3.6)
\]
We can also write (3.5) in a “programming language” format, expressing it as a set of instructions for computing MAJ given the basic operations AND, OR, NOT:

```python
def MAJ(X[0],X[1],X[2]):
    firstpair = AND(X[0],X[1])
    secondpair = AND(X[1],X[2])
    thirdpair = AND(X[0],X[2])
    temp = OR(secondpair,thirdpair)
    return OR(firstpair,temp)
```

### 3.2.1 Some properties of AND and OR
Like standard addition and multiplication, the functions AND and OR satisfy the properties of commutativity: \( a \lor b = b \lor a \) and \( a \land b = b \land a \) and associativity: \( (a \lor b) \lor c = a \lor (b \lor c) \) and \( (a \land b) \land c = a \land (b \land c) \). As in the case of addition and multiplication, we often drop the parenthesis and write \( a \lor b \lor c \lor d \) for \( ((a \lor b) \lor c) \lor d \), and similarly OR’s and AND’s of more terms. They also satisfy a variant of the distributive law:

**Solved Exercise 3.1 — Distributive law for AND and OR.** Prove that for every \( a, b, c \in \{0, 1\} \),
\[
(a \land (b \lor c)) = (a \land b) \lor (a \land c).
\]

**Solution:**
We can prove this by enumerating over all the 8 possible values for \( a, b, c \in \{0, 1\} \) but it also follows from the standard distributive law. Suppose that we identify any positive integer with “true” and the value zero with “false”. Then for every numbers \( u, v \in \mathbb{N} \), \( u + v \) is positive if and only if \( u \lor v \) is true and \( u \cdot v \) is positive if and only if \( u \land v \) is true. This means that for every \( a, b, c \in \{0, 1\} \), the expression \( a \land (b \lor c) \) is true if and only if \( a \cdot (b + c) \) is positive, and the expression \( (a \land b) \lor (a \land c) \) is true if and only if \( a \cdot (b + c) \) is positive, but by the standard distributive law \( a \cdot (b + c) = a \cdot b + a \cdot c \) and hence the former expression is true if and only if the latter one is.

### 3.2.2 Extended example: Computing XOR from AND, OR, and NOT
Let us see how we can obtain a different function from the same building blocks. Define XOR : \( \{0, 1\}^2 \to \{0, 1\} \) to be the function \( XOR(a, b) = a + b \mod 2 \). That is, \( XOR(0, 0) = XOR(1, 1) = 0 \) and \( XOR(1, 0) = XOR(0, 1) = 1 \). We claim that we can construct XOR using only AND, OR, and NOT.
As usual, it is a good exercise to try to work out the algorithm for XOR using AND, OR and NOT on your own before reading further.

The following algorithm computes XOR using AND, OR, and NOT:

Algorithm 3.2 — XOR from AND/OR/NOT.

Input: \(a, b \in \{0, 1\}\).
Output: XOR\((a, b)\)
1: \(w_1 \leftarrow AND(a, b)\)
2: \(w_2 \leftarrow NOT(w_1)\)
3: \(w_3 \leftarrow OR(a, b)\)
4: return \(AND(w_2, w_3)\)

Lemma 3.3 For every \(a, b \in \{0, 1\}\), on input \(a, b\), Algorithm 3.2 outputs \(a + b \mod 2\).

Proof. For every \(a, b\), XOR\((a, b)\) = 1 if and only if \(a\) is different from \(b\). On input \(a, b \in \{0, 1\}\), Algorithm 3.2 outputs AND\((w_2, w_3)\) where \(w_2 = NOT(AND(a, b))\) and \(w_3 = OR(a, b)\).

• If \(a = b = 0\) then \(w_3 = OR(a, b) = 0\) and so the output will be 0.
• If \(a = b = 1\) then \(AND(a, b) = 1\) and so \(w_2 = NOT(AND(a, b)) = 0\) and the output will be 0.
• If \(a = 1\) and \(b = 0\) (or vice versa) then both \(w_3 = OR(a, b) = 1\) and \(w_1 = AND(a, b) = 0\), in which case the algorithm will output \(OR(NOT(w_1), w_3) = 1\).

We can also express Algorithm 3.2 using a programming language. Specifically, the following is a Python program that computes the XOR function:

```python
def AND(a,b): return a*b
def OR(a,b): return 1-(1-a)*(1-b)
def NOT(a): return 1-a

def XOR(a,b):
    w1 = AND(a,b)
    w2 = NOT(w1)
    w3 = OR(a,b)
    return AND(w2,w3)
```

# Test out the code
Solved Exercise 3.2 — Compute $XOR$ on three bits of input. Let $XOR_3 : \{0,1\}^3 \to \{0,1\}$ be the function defined as $XOR_3(a,b,c) = a + b + c \mod 2$. That is, $XOR_3(a,b,c) = 1$ if $a+b+c$ is odd, and $XOR_3(a,b,c) = 0$ otherwise. Show that you can compute $XOR_3$ using AND, OR, and NOT. You can express it as a formula, use a programming language such as Python, or use a Boolean circuit.

Solution:

Addition modulo two satisfies the same properties of associativity $((a + b) + c = (a + b) + c)$ and commutativity $(a + b = b + a)$ as standard addition. This means that, if we define $a \oplus b$ to equal $a + b \mod 2$, then

$$XOR_3(a,b,c) = (a \oplus b) \oplus c$$

(3.7)

or in other words

$$XOR_3(a,b,c) = XOR(XOR(a,b),c).$$

(3.8)

Since we know how to compute $XOR$ using AND, OR, and NOT, we can compose this to compute $XOR_3$ using the same building blocks. In Python this corresponds to the following program:

```python
def XOR3(a,b,c):
    w1 = AND(a,b)
    w2 = NOT(w1)
    w3 = OR(a,b)
    w4 = AND(w2,w3)
    w5 = AND(w4,c)
    w6 = NOT(w5)
    w7 = OR(w4,c)
    return AND(w6,w7)
```

# Let's test this out
print([f"XOR3({a},{b},{c})={XOR3(a,b,c)}" for a in [0,1] for b in [0,1] for c in [0,1]])
```
Try to generalize the above examples to obtain a way to compute $\text{XOR}_n : \{0, 1\}^n \rightarrow \{0, 1\}$ for every $n$ using at most $4n$ basic steps involving applications of a function in $\{\text{AND, OR, NOT}\}$ to outputs or previously computed values.

3.2.3 Informally defining “basic operations” and “algorithms”

We have seen that we can obtain at least some examples of interesting functions by composing together applications of AND, OR, and NOT. This suggests that we can use AND, OR, and NOT as our “basic operations”, hence obtaining the following definition of an “algorithm”:

**Semi-formal definition of an algorithm:** An algorithm consists of a sequence of steps of the form “compute a new value by applying AND, OR, or NOT to previously computed values”.

An algorithm $A$ computes a function $F$ if for every input $x$ to $F$, if we feed $x$ as input to the algorithm, the value computed in its last step is $F(x)$.

There are several concerns that are raised by this definition:

1. First and foremost, this definition is indeed too informal. We do not specify exactly what each step does, nor what it means to “feed $x$ as input”.

2. Second, the choice of AND, OR or NOT seems rather arbitrary. Why not XOR and MAJ? Why not allow operations like addition and multiplication? What about any other logical constructions such if/then or while?

3. Third, do we even know that this definition has anything to do with actual computing? If someone gave us a description of such an algorithm, could we use it to actually compute the function in the real world?

These concerns will to a large extent guide us in the upcoming chapters. Thus you would be well advised to re-read the above informal definition and see what you think about these issues.

A large part of this book will be devoted to addressing the above issues. We will see that:
1. We can make the definition of an algorithm fully formal, and so give a precise mathematical meaning to statements such as “Algorithm \( A \) computes function \( f \).

2. While the choice of \( \text{AND}/\text{OR}/\text{NOT} \) is arbitrary, and we could just as well chose some other functions, we will also see this choice does not matter much. We will see that we would obtain the same computational power if we used instead for addition and multiplication, and essentially every other operation that could be reasonably thought of as a basic step.

3. It turns out that we can and do compute such \( \text{AND}/\text{OR}/\text{NOT} \) based algorithms” in the real world. First of all, such an algorithm is clearly well specified, and so can be executed by a human with a pen and paper. Second, there are a variety of ways to mechanize this computation. We’ve already seen that we can write Python code that corresponds to following such a list of instructions. But in fact we can directly implement operations such as \( \text{AND}, \text{OR}, \) and \( \text{NOT} \) via electronic signals using components known as transistors. This is how modern electronic computers operate.

In the remainder of this chapter, and the rest of this book, we will begin to answer some of these questions. We will see more examples of the power of simple operations to compute more complex operations including addition, multiplication, sorting and more. We will also discuss how to physically implement simple operations such as \( \text{AND}, \text{OR} \) and \( \text{NOT} \) using a variety of technologies.

3.3 BOOLEAN CIRCUITS

Boolean circuits provide a precise notion of “composing basic operations together”. A Boolean circuit (see Fig. 3.9) is composed of gates and inputs that are connected by wires. The wires carry a signal that represents either the value 0 or 1. Each gate corresponds to either the \( \text{OR}, \text{AND}, \) or \( \text{NOT} \) operation. An \( \text{OR} \) gate has two incoming wires, and one or more outgoing wires. If these two incoming wires carry the signals \( a \) and \( b \) (for \( a, b \in \{0, 1\} \)), then the signal on the outgoing wires will be \( \text{OR}(a, b) \). \( \text{AND} \) and \( \text{NOT} \) gates are defined similarly. The inputs have only outgoing wires. If we set a certain input to a value \( a \in \{0, 1\} \), then this value is propagated on all the wires outgoing from it. We also designate some gates as output gates, and their value corresponds to the result of evaluating the circuit. For example, Fig. 3.8 gives such a circuit for the \( \text{XOR} \) function, following Section 3.2.2. We evaluate an \( n \)-input Boolean circuit \( C \) on an input \( x \in \{0, 1\}^n \) by placing the bits of \( x \) on the inputs, and then propagating the values on the wires until we reach an output, see Fig. 3.9.
**Remark 3.4 — Physical realization of Boolean circuits.**

Boolean circuits are a mathematical model that does not necessarily correspond to a physical object, but they can be implemented physically. In physical implementation of circuits, the signal is often implemented by electric potential or voltage on a wire, where for example voltage above a certain level is interpreted as a logical value of 1, and below a certain level is interpreted as a logical value of 0. Section 3.4 discusses physical implementation of Boolean circuits (with examples including using electrical signals such as in silicon-based circuits, but also biological and mechanical implementations as well).

---

**Figure 3.9:** A Boolean Circuit consists of gates that are connected by wires to one another and the inputs. The left side depicts a circuit with 2 inputs and 5 gates, one of which is designated the output gate. The right side depicts the evaluation of this circuit on the input $x \in \{0, 1\}^2$ with $x_0 = 1$ and $x_1 = 0$. The value of every gate is obtained by applying the corresponding function (AND, OR, or NOT) to values on the wire(s) that enter it. The output of the circuit on a given input is the value of the output gate(s). In this case, the circuit computes the XOR function and hence it outputs 1 on the input 10.

---

**Solved Exercise 3.3 — All equal function.** Define $ALLEQ : \{0, 1\}^4 \to \{0, 1\}$ to be the function that on input $x \in \{0, 1\}^4$ outputs 1 if and only if $x_0 = x_1 = x_2 = x_3$. Give a Boolean circuit for computing $ALLEQ$.

**Solution:**

Another way to describe the function $ALLEQ$ is that it outputs 1 on an input $x \in \{0, 1\}^4$ if and only if $x = 0^4$ or $x = 1^4$. We can phrase the condition $x = 1^4$ as $x_0 \land x_1 \land x_2 \land x_3$ which can be computed using three AND gates. Similarly we can phrase the condition $x = 0^4$ as $\overline{x_0} \land \overline{x_1} \land \overline{x_2} \land \overline{x_3}$ which can be computed using four NOT gates and three AND gates. The output of $ALLEQ$ is the OR of these two conditions, which results in the circuit of 4 NOT gates, 6 AND gates, and one OR gate presented in Fig. 3.10.

---

**3.3.1 Boolean circuits: a formal definition**

We defined Boolean circuits informally as obtained by connecting AND, OR, and NOT gates via wires so as to produce an output from an input. However, to be able to prove theorems about the existence or...
non-existence of Boolean circuits for computing various functions we need to:

1. Formally define a Boolean circuit as a mathematical object.

2. Formally define what it means for a circuit $C$ to compute a function $f$.

We now proceed to do so. We will define a Boolean circuit as a labeled Directed Acyclic Graph (DAG). The vertices of the graph correspond to the gates and inputs of the circuit, and the edges of the graph correspond to the wires. A wire from an input or gate $u$ to a gate $v$ in the circuit corresponds to a directed edge between the corresponding vertices. The inputs are vertices with no incoming edges, while each gate has the appropriate number of incoming edges based on the function it computes. (That is, AND and OR gates have two in-neighbors, while NOT gates have one in-neighbor.) The formal definition is as follows (see also Fig. 3.11):

**Definition 3.5 — Boolean Circuits.** Let $n, m, s$ be positive integers with $s \geq m$. A Boolean circuit with $n$ inputs, $m$ outputs, and $s$ gates, is a labeled directed acyclic graph (DAG) $G = (V, E)$ with $s + n$ vertices satisfying the following properties:

- Exactly $n$ of the vertices have no in-neighbors. These vertices are known as inputs and are labeled with the $n$ labels $X[0], ..., X[n - 1]$. Each input has at least one out-neighbor.

**Figure 3.11:** A Boolean Circuit is a labeled directed acyclic graph (DAG). It has $n$ input vertices, which are marked with $X[0], ..., X[n - 1]$ and have no incoming edges, and the rest of the vertices are gates. An AND, OR, or NOT gate has two or one incoming edges. If the circuit has $m$ outputs, then $m$ of the gates are known as outputs and are marked with $Y[0], ..., Y[m - 1]$. When we evaluate a circuit $C$ on an input $x \in \{0, 1\}^n$, we start by setting the value of the input vertices to $x_0, ..., x_{n - 1}$ and then propagate the values, assigning to each gate $g$ the result of applying the operation of $g$ to the values of $g$'s in-neighbors. The output of the circuit is the value assigned to the output gates.
The other $s$ vertices are known as gates. Each gate is labeled with $\land$, $\lor$ or $\lnot$. Gates labeled with $\land$ (AND) or $\lor$ (OR) have two in-neighbors. Gates labeled with $\lnot$ (NOT) have one in-neighbor. We will allow parallel edges. ¹

Exactly $m$ of the gates are also labeled with the $m$ labels $Y[0], \ldots, Y[m-1]$ (in addition to their label $\land/\lor/\lnot$). These are known as outputs.

The size of a Boolean circuit is the number of gates it contains.

¹ Having parallel edges means an AND or OR gate $u$ can have both its in-neighbors be the same gate $v$. Since $\text{AND}(a, a) = \text{OR}(a, a) = a$ for every $a \in \{0, 1\}$, such parallel edges don’t help in computing new values in circuits with AND/OR/NOT gates. However, we will see circuits with more general sets of gates later on.

---

If $C$ is a circuit with $n$ inputs and $m$ outputs, and $x \in \{0, 1\}^n$, then we can compute the output of $C$ on the input $x$ in the natural way: assign the input vertices $X[0], \ldots, X[n-1]$ the values $x_0, \ldots, x_{n-1}$, apply each gate on the values of its in-neighbors, and then output the values that correspond to the output vertices. Formally, this is defined as follows:

**Definition 3.6 — Computing a function via a Boolean circuit.** Let $C$ be a Boolean circuit with $n$ inputs and $m$ outputs. For every $x \in \{0, 1\}^n$, the output of $C$ on the input $x$, denoted by $C(x)$, is defined as the result of the following process:

We let $h : V \to \mathbb{N}$ be the minimal layering of $C$ (aka topological sorting, see Theorem 1.26). We let $L$ be the maximum layer of $h$, and for $\ell = 0, 1, \ldots, L$ we do the following:

- For every $v$ in the $\ell$-th layer (i.e., $v$ such that $h(v) = \ell$) do:
  - If $v$ is an input vertex labeled with $X[i]$ for some $i \in [n]$, then we assign to $v$ the value $x_i$.
  - If $v$ is a gate vertex labeled with $\land$ and with two in-neighbors $u, w$ then we assign to $v$ the AND of the values assigned to $u$ and $w$. (Since $u$ and $w$ are in-neighbors of $v$, they are in lower layer than $v$, and hence their values have already been assigned.)
– If \( v \) is a gate vertex labeled with \( \lor \) and with two in-neighbors \( u, w \) then we assign to \( v \) the OR of the values assigned to \( u \) and \( w \).

– If \( v \) is a gate vertex labeled with \( \neg \) and with one in-neighbor \( u \) then we assign to \( v \) the negation of the value assigned to \( u \).

• The result of this process is the value \( y \in \{0, 1\}^m \) such that for every \( j \in [m] \), \( y_j \) is the value assigned to the vertex with label \( Y[j] \).

Let \( f : \{0, 1\}^n \to \{0, 1\}^m \). We say that the circuit \( C \) computes \( f \) if for every \( x \in \{0, 1\}^n \), \( C(x) = f(x) \).

**Remark 3.7 — Boolean circuits nitpicks (optional).** In phrasing **Definition 3.5**, we’ve made some technical choices that are not very important, but will be convenient for us later on. Having parallel edges means an AND or OR gate \( u \) can have both its in-neighbors be the same gate \( v \). Since \( AND(a, a) = OR(a, a) = a \) for every \( a \in \{0, 1\} \), such parallel edges don’t help in computing new values in circuits with AND/OR/NOT gates. However, we will see circuits with more general sets of gates later on. The condition that every input vertex has at least one out-neighbor is also not very important because we can always add “dummy gates” that touch these inputs. However, it is convenient since it guarantees that (since every gate has at most two in-neighbors) that the number of inputs in a circuit is never larger than twice its size.

### 3.3.2 Equivalence of circuits and straight-line programs

We have seen two ways to describe how to compute a function \( f \) using \( AND, OR \) and \( NOT \):

• A **Boolean circuit**, defined in **Definition 3.5**, computes \( f \) by connecting via wires \( AND, OR \), and \( NOT \) gates to the inputs.

• We can also describe such a computation using a **straight-line program** that has lines of the form \( foo = AND(bar, blah) \), \( foo = OR(bar, blah) \) and \( foo = NOT(bar) \) where \( foo, bar \) and \( blah \) are variable names. (We call this a **straight-line program** since it contains no loops or branching (e.g., if/then) statements.)

We now formally define the AON-CIRC programming language (“AON” stands for \( AND/OR/NOT \)) which has the above operations, and show that it is equivalent to Boolean circuits.
We follow the common programming languages convention of using names such as foo, bar, baz, blah as stand-ins for generic identifiers. A variable identifier in our programming language can be any combination of letters, numbers, underscores, and brackets. The appendix contains a full formal specification of our programming language.

Definition 3.8 — AON-CIRC Programming language. An AON-CIRC program is a string of lines of the form foo = AND(bar, blah), foo = OR(bar, blah) and foo = NOT(bar) where foo, bar and blah are variable names. Variables of the form X[i] are known as input variables, and variables of the form Y[j] are known as output variables. In every line, the variables on the right hand side of the assignment operators must either be input variables or variables that have already been assigned a value before.

A valid AON-CIRC program P includes input variables of the form X[0]...,X[n − 1] and output variables of the form Y[0],...,Y[m − 1] for some n, m ≥ 1. If P is valid AON-CIRC program and x ∈ {0, 1}^n, then we define the output of P on input x, denoted by P(x), to be the string y ∈ {0, 1}^m corresponding to the values of the output variables Y[0],...,Y[m − 1] in the execution of P where we initialize the input variables X[0],...,X[n − 1] to the values x_0,...,x_{n−1}.

We say that such an AON-CIRC program P computes a function f : {0, 1}^n → {0, 1}^m if P(x) = f(x) for every x ∈ {0, 1}^n.

AON-CIRC is not a practical programming language: it was designed for pedagogical purposes only, as a way to model computation as composition of AND, OR, and NOT. However, AON-CIRC can still be easily implemented on a computer. The following solved exercise gives an example of an AON-CIRC program.

Solved Exercise 3.4 Consider the following function CMP : {0, 1}^4 → {0, 1} that on four input bits a, b, c, d ∈ {0, 1}, outputs 1 iff the number represented by (a, b) is larger than the number represented by (c, d). That is CMP(a, b, c, d) = 1 iff 2a + b > 2c + d.

Write an AON-CIRC program to compute CMP.

Solution:

Writing such a program is tedious but not truly hard. To compare two numbers we first compare their most significant digit, and then go down to the next digit and so on and so forth. In this case where the numbers have just two binary digits, these comparisons are particularly simple. The number represented by (a, b) is larger than the number represented by (c, d) if and only if one of the following conditions happens:

1. The most significant bit a of (a, b) is larger than the most significant bit c of (c, d).
or

2. The two most significant bits \(a\) and \(c\) are equal, but \(b > d\).

   Another way to express the same condition is the following: the number \((a, b)\) is larger than \((c, d)\) iff \(a > c\) OR \((a \geq c\) AND \(b > d\)).

   For binary digits \(\alpha, \beta\), the condition \(\alpha > \beta\) is simply that \(\alpha = 1\) and \(\beta = 0\) or \(AND(\alpha, NOT(\beta)) = 1\), and the condition \(\alpha \geq \beta\) is simply \(OR(\alpha, NOT(\beta)) = 1\). Together these observations can be used to give the following AON-CIRC program to compute \(CMP\):

   \[
   \begin{align*}
   \text{temp}_1 &= \text{NOT}(X[2]) \\
   \text{temp}_2 &= \text{AND}(X[0], \text{temp}_1) \\
   \text{temp}_3 &= \text{OR}(X[0], \text{temp}_1) \\
   \text{temp}_4 &= \text{NOT}(X[3]) \\
   \text{temp}_5 &= \text{AND}(X[1], \text{temp}_4) \\
   \text{temp}_6 &= \text{AND}(\text{temp}_5, \text{temp}_3) \\
   Y[0] &= \text{OR}(\text{temp}_2, \text{temp}_6)
   \end{align*}
   \]

   We can also present this 8-line program as a circuit with 8 gates, see Fig. 3.12.

It turns out that AON-CIRC programs and Boolean circuits have exactly the same power:

**Theorem 3.9 — Equivalence of circuits and straight-line programs.** Let \(f : \{0, 1\}^n \rightarrow \{0, 1\}^m\) and \(s \geq m\) be some number. Then \(f\) is computable by a Boolean circuit of \(s\) gates if and only if \(f\) is computable by an AON-CIRC program of \(s\) lines.

**Proof Idea:**

The idea is simple - AON-CIRC programs and Boolean circuits are just different ways of describing the exact same computational process. For example, an \(AND\) gate in a Boolean circuit corresponding to computing the \(AND\) of two previously-computed values. In a AON-CIRC program this will correspond to the line that stores in a variable the \(AND\) of two previously-computed variables.

\[\star\]

This proof of Theorem 3.9 is simple at heart, but all the details it contains can make it a little cumbersome to read. You might be better off trying to work it out yourself before reading it. Our GitHub repository contains a “proof by Python” of Theorem 3.9: implementation of functions circuit2prog and prog2circuits.

---

**Figure 3.12:** A circuit for computing the CMP function. The evaluation of this circuit on \((1, 1, 1, 0)\) yields the output 1, since the number 3 (represented in binary as 11) is larger than the number 2 (represented in binary as 10).
Proof of Theorem 3.9. Let \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \). Since the theorem is an “if and only if” statement, to prove it we need to show both directions: translating an AON-CIRC program that computes \( f \) into a circuit that computes \( f \), and translating a circuit that computes \( f \) into an AON-CIRC program that does so.

We start with the first direction. Let \( P \) be an \( s \) line AON-CIRC that computes \( f \). We define a circuit \( C \) as follows: the circuit will have \( n \) inputs and \( s \) gates. For every \( i \in [s] \), if the \( i \)-th line has the form \( \text{foo} = \text{AND} (\text{bar}, \text{blah}) \) then the \( i \)-th gate in the circuit will be an AND gate that is connected to gates \( j \) and \( k \) where \( j \) and \( k \) correspond to the last lines before \( i \) where the variables \( \text{bar} \) and \( \text{blah} \) (respectively) were written to. (For example, if \( i = 57 \) and the last line bar was written to is 35 and the last line blah was written to is 17 then the two in-neighbors of gate 57 will be gates 35 and 17.) If either bar or blah is an input variable then we connect the gate to the corresponding input vertex instead. If \( \text{foo} \) is an output variable of the form \( \text{Y}[\cdot] \) then we add the same label to the corresponding gate to mark it as an output gate. We do the analogous operations if the \( i \)-th line involves an OR or a NOT operation (except that we use the corresponding OR or NOT gate, and in the latter case have only one in-neighbor instead of two).

For every input \( x \in \{0, 1\}^n \), if we run the program \( P \) on \( x \), then the value written that is computed in the \( i \)-th line is exactly the value that will be assigned to the \( i \)-th gate if we evaluate the circuit \( C \) on \( x \). Hence \( C(x) = P(x) \) for every \( x \in \{0, 1\}^n \).

For the other direction, let \( C \) be a circuit of \( s \) gates and \( n \) inputs that computes the function \( f \). We sort the gates according to a topological order and write them as \( v_0, \ldots, v_{s-1} \). We now can create a program \( P \) of \( s \) lines as follows. For every \( i \in [s] \), if \( v_i \) is an AND gate with in-neighbors \( v_j, v_k \) then we will add a line to \( P \) of the form \( \text{temp}_i = \text{AND} (\text{temp}_j, \text{temp}_k) \), unless one of the vertices is an input vertex or an output gate, in which case we change this to the form \( \text{X}[\cdot] \) or \( \text{Y}[\cdot] \) appropriately. Because we work in topological ordering, we are guaranteed that the in-neighbors \( v_j \) and \( v_k \) correspond to variables that have already been assigned a value. We do the same for OR and NOT gate. Once again, one can verify that for every input \( x \), the value \( P(x) \) will equal \( C(x) \) and hence the program computes the same function as the circuit. (Note that since \( C \) is a valid circuit per Definition 3.5, every input vertex of \( C \) has at least one out-neighbor and there are exactly \( m \) output gates labeled 0, \ldots, \( m - 1 \); hence all the variables \( X[0], \ldots, X[n-1] \) and \( Y[0], \ldots, Y[m-1] \) will appear in the program \( P \).)
3.4 PHYSICAL IMPLEMENTATIONS OF COMPUTING DEVICES (DIGRESSION)

Computation is an abstract notion that is distinct from its physical implementations. While most modern computing devices are obtained by mapping logical gates to semi-conductor based transistors, over history people have computed using a huge variety of mechanisms, including mechanical systems, gas and liquid (known as fluidics), biological and chemical processes, and even living creatures (e.g., see Fig. 3.14 or this video for how crabs or slime mold can be used to do computations).

In this section we will review some of these implementations, both so you can get an appreciation of how it is possible to directly translate Boolean circuits to the physical world, without going through the entire stack of architecture, operating systems, and compilers, as well as to emphasize that silicon-based processors are by no means the only way to perform computation. Indeed, as we will see in Chapter 22, a very exciting recent line of works involves using different media for computation that would allow us to take advantage of quantum mechanical effects to enable different types of algorithms.

3.4.1 Transistors

A transistor can be thought of as an electric circuit with two inputs, known as source and gate and an output, known as the sink. The gate controls whether current flows from the source to the sink. In a standard transistor, if the gate is “ON” then current can flow from the source to the sink and if it is “OFF” then it can’t. In a complementary transistor this is reversed: if the gate is “OFF” then current can flow from the source to the sink and if it is “ON” then it can’t.

There are several ways to implement the logic of a transistor. For example, we can use faucets to implement it using water pressure (e.g. Fig. 3.15). This might seem as merely a curiosity but there is a field known as fluidics concerned with implementing logical operations using liquids or gasses. Some of the motivations include operating in extreme environmental conditions such as in space or a battlefield, where standard electronic equipment would not survive.

The standard implementations of transistors use electrical current. One of the original implementations used vacuum tubes. As its name implies, a vacuum tube is a tube containing nothing (i.e., a vacuum) and where a priori electrons could freely flow from source (a wire) to the sink (a plate). However, there is a gate (a grid) between the two, where modulating its voltage can block the flow of electrons.
Early vacuum tubes were roughly the size of lightbulbs (and looked very much like them too). In the 1950’s they were supplanted by transistors, which implement the same logic using semiconductors which are materials that normally do not conduct electricity but whose conductivity can be modified and controlled by inserting impurities (“doping”) and an external electric field (this is known as the field effect). In the 1960’s computers were started to be implemented using integrated circuits which enabled much greater density. In 1965, Gordon Moore predicted that the number of transistors per integrated circuit would double every year (see Fig. 3.16), and that this would lead to “such wonders as home computers —or at least terminals connected to a central computer— automatic controls for automobiles, and personal portable communications equipment”. Since then, (adjusted versions of) this so-called “Moore’s law” have been running strong, though exponential growth cannot be sustained forever, and some physical limitations are already becoming apparent.

3.4.2 Logical gates from transistors
We can use transistors to implement various Boolean functions such as AND, OR, and NOT. For each a two-input gate \( G : \{0, 1\}^2 \to \{0, 1\} \), such an implementation would be a system with two input wires \( x, y \) and one output wire \( z \), such that if we identify high voltage with “1” and low voltage with “0”, then the wire \( z \) will equal to “1” if and only if applying \( G \) to the values of the wires \( x \) and \( y \) is 1 (see Fig. 3.19 and Fig. 3.20). This means that there exists a AND/OR/NOT circuit to compute a function \( g : \{0, 1\}^n \to \{0, 1\}^m \), then we can compute \( g \) in the physical world using transistors as well.

3.4.3 Biological computing
Computation can be based on biological or chemical systems. For example the lac operon produces the enzymes needed to digest lactose only if the conditions \( x \land (\neg y) \) hold where \( x \) is “lactose is present” and \( y \) is “glucose is present”. Researchers have managed to create transistors, and from them logic gates, based on DNA molecules (see also Fig. 3.21). One motivation for DNA computing is to achieve increased parallelism or storage density; another is to create “smart biological agents” that could perhaps be injected into bodies, replicate themselves, and fix or kill cells that were damaged by a disease such as cancer. Computing in biological systems is not restricted of course to DNA. Even larger systems such as flocks of birds can be considered as computational processes.

3.4.4 Cellular automata and the game of life
Cellular automata is a model of a system composed of a sequence of cells, which of which can have a finite state. At each step, a cell up-
dates its state based on the states of its neighboring cells and some simple rules. As we will discuss later in this book (see Section 7.4), cellular automata such as Conway’s “Game of Life” can be used to simulate computation gates.

### 3.4.5 Neural networks

One computation device that we all carry with us is our own brain. Brains have served humanity throughout history, doing computations that range from distinguishing prey from predators, through making scientific discoveries and artistic masterpieces, to composing witty 280 character messages. The exact working of the brain is still not fully understood, but one common mathematical model for it is a (very large) neural network.

A neural network can be thought of as a Boolean circuit that instead of AND/OR/NOT uses some other gates as the basic basis. For example, one particular basis we can use are threshold gates. For every vector \( w = (w_0, \ldots, w_{k-1}) \) of integers and integer \( t \) (some or all of whom could be negative), the threshold function corresponding to \( w, t \) is the function \( T_{w,t} : \{0, 1\}^k \to \{0, 1\} \) that maps \( x \in \{0, 1\}^k \) to 1 if and only if \( \sum_{i=0}^{k-1} w_i x_i \geq t \). For example, the threshold function \( T_{w,t} \) corresponding to \( w = (1, 1, 1, 1, 1) \) and \( t = 3 \) is simply the majority function \( \text{MAJ}_5 \) on \( \{0, 1\}^5 \). Threshold gates can be thought of as an approximation for neuron cells that make up the core of human and animal brains. To a first approximation, a neuron has \( k \) inputs and a single output and the neurons “fires” or “turns on” its output when those signals pass some threshold.

Many machine learning algorithms use artificial neural networks whose purpose is not to imitate biology but rather to perform some computational tasks, and hence are not restricted to threshold or other biologically-inspired gates. Generally, a neural network is often described as operating on signals that are real numbers, rather than 0/1 values, and where the output of a gate on inputs \( x_0, \ldots, x_{k-1} \) is obtained by applying \( f(\sum_i w_i x_i) \) where \( f : \mathbb{R} \to \mathbb{R} \) is an an activation function such as rectified linear unit (ReLU), Sigmoid, or many others (see Fig. 3.23). However, for the purposes of our discussion, all of the above are equivalent (see also Exercise 3.13). In particular we can reduce the setting of real inputs to binary inputs by representing a real number in the binary basis, and multiplying the weight of the bit corresponding to the \( i^{th} \) digit by \( 2^i \).

### 3.4.6 A computer made from marbles and pipes

We can implement computation using many other physical media, without any electronic, biological, or chemical components. Many suggestions for mechanical computers have been put forward, going
back at least to Gottfried Leibniz’ computing machines from the 1670s and Charles Babbage’s 1837 plan for a mechanical “Analytical Engine”. As one example, Fig. 3.24 shows a simple implementation of a NAND (negation of AND, see Section 3.5) gate using marbles going through pipes. We represent a logical value in \( \{0, 1\} \) by a pair of pipes, such that for every \( a, b \in \{0, 1\} \), if two marbles are rolling toward the object in the \( a \)-pipe of the first pair and the \( b \)-pipe of the second pair, then a marble will roll out of the object in the \( \text{NAND}(a, b) \)-pipe of the outgoing pair. In fact, there is even a commercially-available educational game that uses marbles as a basis of computing, see Fig. 3.26.

### 3.5 THE NAND FUNCTION

The NAND function is another simple function that is extremely useful for defining computation. It is the function mapping \( \{0, 1\}^2 \) to \( \{0, 1\} \) defined by:

\[
\text{NAND}(a, b) = \begin{cases} 
0 & a = b = 1 \\
1 & \text{otherwise}
\end{cases}
\] (3.9)

As its name implies, NAND is the NOT of AND (i.e., \( \text{NAND}(a, b) = \text{NOT}(\text{AND}(a, b)) \)), and so we can clearly compute NAND using AND and NOT. Interestingly, the opposite direction holds as well:

**Theorem 3.10 — NAND computes AND, OR, NOT.** We can compute AND, OR, and NOT by composing only the NAND function.

**Proof.** We start with the following observation. For every \( a \in \{0, 1\} \), \( \text{AND}(a, a) = a \). Hence, \( \text{NAND}(a, a) = \text{NOT}(\text{AND}(a, a)) = \text{NOT}(a) \). This means that NAND can compute NOT. By the principle of “double negation”, \( \text{AND}(a, b) = \text{NOT}(\text{NOT}(\text{AND}(a, b))) \), and hence we can use NAND to compute AND as well. Once we can compute AND and NOT, we can compute OR using “De Morgan’s Law”: \( \text{OR}(a, b) = \text{NOT}(\text{AND}(\text{NOT}(a), \text{NOT}(b))) \) (which can also be written as \( a \lor b = \overline{a} \land \overline{b} \)) for every \( a, b \in \{0, 1\} \).

> Theorem 3.10’s proof is very simple, but you should make sure that (i) you understand the statement of
We can use NAND to compute many other functions, as demonstrated in the following exercise.

**Solved Exercise 3.5 — Compute majority with NAND.** Let \( \text{MAJ} : \{0, 1\}^3 \to \{0, 1\} \) be the function that on input \( a, b, c \) outputs 1 iff \( a + b + c \geq 2 \). Show how to compute \( \text{MAJ} \) using a composition of NAND’s.

**Solution:**
Recall that (3.5) states that

\[
\text{MAJ}(x_0, x_1, x_2) = \text{OR}(\text{AND}(x_0, x_1), \text{OR}(\text{AND}(x_1, x_2), \text{AND}(x_0, x_2))).
\]

We can use Theorem 3.10 to replace all the occurrences of AND and OR with NAND’s. Specifically, we can use the equivalence
\[
\text{AND}(a, b) = \text{NOT}(\text{NAND}(a, b)), \text{OR}(a, b) = \text{NAND}(\text{NOT}(a), \text{NOT}(b)),
\]
and \( \text{NOT}(a) = \text{NAND}(a, a) \) to replace the righthand side of (3.10) with an expression involving only NAND, yielding that \( \text{MAJ}(a, b, c) \) is equivalent the (somewhat unwieldy) expression

\[
\text{NAND}
\left(
\text{NAND}
\left(
\text{NAND}(\text{NAND}(a, b), \text{NAND}(a, c)),
\text{NAND}(\text{NAND}(a, b), \text{NAND}(a, c))
\right),
\text{NAND}(b, c)
\right).
\]

The same formula can also be expressed as a circuit with NAND gates, see Fig. 3.27.

### 3.5.1 NAND Circuits

We define **NAND Circuits** as circuits in which all the gates are NAND operations. Such a circuit again corresponds to a directed acyclic graph (DAG) since all the gates correspond to the same function (i.e., NAND), we do not even need to label them, and all gates have in-degree exactly two. Despite their simplicity, NAND circuits can be quite powerful.
Example 3.11 — NAND circuit for XOR. Recall the XOR function which maps $x_0, x_1 \in \{0, 1\}$ to $x_0 + x_1 \mod 2$. We have seen in Section 3.2.2 that we can compute XOR using AND, OR, and NOT, and so by Theorem 3.10 we can compute it using only NAND’s. However, the following is a direct construction of computing XOR by a sequence of NAND operations:

1. Let $u = \text{NAND}(x_0, x_1)$.
2. Let $v = \text{NAND}(x_0, u)$
3. Let $w = \text{NAND}(x_1, u)$.
4. The XOR of $x_0$ and $x_1$ is $y_0 = \text{NAND}(v, w)$.

One can verify that this algorithm does indeed compute XOR by enumerating all the four choices for $x_0, x_1 \in \{0, 1\}$. We can also represent this algorithm graphically as a circuit, see Fig. 3.28.

In fact, we can show the following theorem:

Theorem 3.12 — NAND is a universal operation. For every Boolean circuit $C$ of $s$ gates, there exists a NAND circuit $C'$ of at most $3s$ gates that computes the same function as $C$.

Proof Idea:

The idea of the proof is to just replace every AND, OR and NOT gate with their NAND implementation following the proof of Theorem 3.10.

Proof of Theorem 3.12. If $C$ is a Boolean circuit, then since, as we’ve seen in the proof of Theorem 3.10, for every $a, b \in \{0, 1\}$

- NOT($a$) = NAND($a, a$)
- AND($a, b$) = NAND(NAND($a, b$), NAND($a, b$))
- OR($a, b$) = NAND(NAND($a, a$), NAND($b, b$))

we can replace every gate of $C$ with at most three NAND gates to obtain an equivalent circuit $C'$. The resulting circuit will have at most $3s$ gates.

Big Idea 3 Two models are equivalent in power if they can be used to compute the same set of functions.
3.5.2 More examples of NAND circuits (optional)

Here are some more sophisticated examples of NAND circuits:

Incrementing integers. Consider the task of computing, given as input a string \( x \in \{0, 1\}^n \) that represents a natural number \( X \in \mathbb{N} \), the representation of \( X + 1 \). That is, we want to compute the function \( \text{INC}_n : \{0, 1\}^n \to \{0, 1\}^{n+1} \) such that for every \( x_0, \ldots, x_{n-1} \), \( \text{INC}_n(x) = y \) which satisfies \( \sum_{i=0}^{n} y_i 2^i = \left( \sum_{i=0}^{n-1} x_i 2^i \right) + 1 \). (For simplicity of notation, in this example we use the representation where the least significant digit is first rather than last.)

The increment operation can be very informally described as follows: "Add 1 to the least significant bit and propagate the carry". A little more precisely, in the case of the binary representation, to obtain the increment of \( x \), we scan \( x \) from the least significant bit onwards, and flip all 1’s to 0’s until we encounter a bit equal to 0, in which case we flip it to 1 and stop.

Thus we can compute the increment of \( x_0, \ldots, x_{n-1} \) by doing the following:

```
Algorithm 3.13 — Compute Increment Function.

Input: \( x_0, x_1, \ldots, x_{n-1} \) representing the number \( \sum_{i=0}^{n-1} x_i \cdot 2^i \) # we use LSB-first representation
Output: \( y \in \{0, 1\}^{n+1} \) such that \( \sum_{i=0}^{n} y_i 2^i = \sum_{i=0}^{n-1} x_i 2^i \)

1: Let \( c_0 \leftarrow 1 \) # we pretend we have a "carry" of 1 initially
2: for \( i = 0, \ldots, n-1 \) do
3:   Let \( y_i \leftarrow \text{XOR}(x_i, c_i) \).
4:   if \( c_i = x_i = 1 \) then
5:     \( c_{i+1} \leftarrow 1 \)
6:   else
7:     \( c_{i+1} \leftarrow 0 \)
8: end if
9: end for
10: Let \( y_n \equiv c_n \).
```

Algorithm 3.13 describes precisely how to compute the increment operation, and can be easily transformed into Python code that performs the same computation, but it does not seem to directly yield a NAND circuit to compute this. However, we can transform this algorithm line by line to a NAND circuit. For example, since for every \( a, \text{NAND}(a, \text{NOT}(a)) = 1 \), we can replace the initial statement \( c_0 = 1 \) with \( c_0 = \text{NAND}(x_0, \text{NAND}(x_0, x_0)) \). We already know how to compute XOR using NAND and so we can use this to implement the operation \( y_i \leftarrow \text{XOR}(x_i, c_i) \). Similarly, we can write the "if" statement as saying \( c_{i+1} \leftarrow \text{AND}(c_i, x_i) \), or in other words
\( c_{i+1} \leftarrow \text{NAND}(\text{NAND}(c_i, x_i), \text{NAND}(c_i, x_i)) \). Finally, the assignment \( y_n = c_n \) can be written as \( y_n = \text{NAND}(\text{NAND}(c_n, c_n), \text{NAND}(c_n, c_n)) \).

Combining these observations yields for every \( n \in \mathbb{N} \), a NAND circuit to compute \( \text{INC}_n \). For example, Fig. 3.29 shows how this circuit looks like for \( n = 4 \).

From increment to addition. Once we have the increment operation, we can certainly compute addition by repeatedly incrementing (i.e., compute \( x + y \) by performing \( \text{INC}(x) \) \( y \) times). However, that would be quite inefficient and unnecessary. With the same idea of keeping track of carries we can implement the “grade-school” addition algorithm and compute the function \( \text{ADD}_n : \{0, 1\}^{2n} \rightarrow \{0, 1\}^{n+1} \) that on input \( x \in \{0, 1\}^{2n} \) outputs the binary representation of the sum of the numbers represented by \( x_0, \ldots, x_{n-1} \) and \( x_{n+1}, \ldots, x_n \):

![Algorithm 3.14 — Addition using NAND.](image)

Once again, Algorithm 3.14 can be translated into a NAND circuit. The crucial observation is that the “if/then” statement simply corresponds to \( c_{i+1} \leftarrow \text{MAJ}_3(u_i, v_i, v_i) \) and we have seen in in Solved Exercise 3.5 that the function \( \text{MAJ}_3 : \{0, 1\}^3 \rightarrow \{0, 1\} \) can be computed using NANDs.

#### 3.5.3 The NAND-CIRC Programming language

Just like we did for Boolean circuits, we can define a programming-language analog of NAND circuits. It is even simpler than the AON-CIRC language since we only have a single operation. We define the **NAND-CIRC Programming Language** to be a programming language where every line has the following form:

\[
\text{foo} = \text{NAND}(\text{bar}, \text{blah})
\]

where foo, bar and blah are variable identifiers.
**Example 3.15 — Our first NAND-CIRC program.** Here is an example of a NAND-CIRC program:

\[
\begin{align*}
    u &= \text{NAND}(X[0], X[1]) \\
    v &= \text{NAND}(X[0], u) \\
    w &= \text{NAND}(X[1], u) \\
    Y[0] &= \text{NAND}(v, w)
\end{align*}
\]

Do you know what function this program computes? Hint: you have seen it before.

Formally, just like we did in Definition 3.8 for AON-CIRC, we can define the notion of computation by a NAND-CIRC program in the natural way:

**Definition 3.16 — Computing by a NAND-CIRC program.** Let \( f : \{0, 1\}^n \to \{0, 1\}^m \) be some function, and let \( P \) be a NAND-CIRC program. We say that \( P \) computes the function \( f \) if:

1. \( P \) has \( n \) input variables \( X[0], \ldots, X[n-1] \) and \( m \) output variables \( Y[0], \ldots, Y[m-1] \).
2. For every \( x \in \{0, 1\}^n \), if we execute \( P \) when we assign to \( X[0], \ldots, X[n-1] \) the values \( x_0, \ldots, x_{n-1} \), then at the end of the execution, the output variables \( Y[0], \ldots, Y[m-1] \) have the values \( y_0, \ldots, y_{m-1} \) where \( y = f(x) \).

As before we can show that NAND circuits are equivalent to NAND-CIRC programs (see Fig. 3.30):

**Theorem 3.17 — NAND circuits and straight-line program equivalence.** For every \( f : \{0, 1\}^n \to \{0, 1\}^m \) and \( s \geq m \), \( f \) is computable by a NAND-CIRC program of \( s \) lines if and only if \( f \) is computable by a NAND circuit of \( s \) gates.

We omit the proof of Theorem 3.17 since it follows along exactly the same lines as the equivalence of Boolean circuits and AON-CIRC program (Theorem 3.9). Given Theorem 3.17 and Theorem 3.12, we know that we can translate every \( s \)-line AON-CIRC program \( P \) into an equivalent NAND-CIRC program of at most \( 3s \) lines. In fact, this translation can be easily done by replacing every line of the form

- \( \text{foo} = \text{AND}(\text{bar}, \text{blah}) \)
- \( \text{foo} = \text{OR}(\text{bar}, \text{blah}) \)
- \( \text{foo} = \text{NOT}(\text{bar}) \)
with the equivalent 1-3 lines that use the NAND operation. Our GitHub repository contains a “proof by code”: a simple Python program AON2NAND that transforms an AON-CIRC into an equivalent NAND-CIRC program.

**Remark 3.18 — Is the NAND-CIRC programming language Turing Complete? (optional note).** You might have heard of a term called “Turing Complete” that is sometimes used to describe programming languages. (If you haven’t, feel free to ignore the rest of this remark: we define this term precisely in Chapter 7.) If so, you might wonder if the NAND-CIRC programming language has this property. The answer is no, or perhaps more accurately, the term “Turing Completeness” is not really applicable for the NAND-CIRC programming language. The reason is that, by design, the NAND-CIRC programming language can only compute finite functions \( f : \{0, 1\}^n \to \{0, 1\}^m \) that take a fixed number of input bits and produce a fixed number of outputs bits. The term “Turing Complete” is only applicable to programming languages for infinite functions that can take inputs of arbitrary length. We will come back to this distinction later on in this book.

### 3.6 EQUIVALENCE OF ALL THESE MODELS

If we put together **Theorem 3.9**, **Theorem 3.12**, and **Theorem 3.17**, we obtain the following result:

**Theorem 3.19 — Equivalence between models of finite computation.** For every sufficiently large \( s, n, m \) and \( f : \{0, 1\}^n \to \{0, 1\}^m \), the following conditions are all equivalent to one another:

- \( f \) can be computed by a Boolean circuit (with \( \wedge, \vee, \neg \) gates) of at most \( O(s) \) gates.
- \( f \) can be computed by an AON-CIRC straight-line program of at most \( O(s) \) lines.
- \( f \) can be computed by a NAND circuit of at most \( O(s) \) gates.
- \( f \) can be computed by a NAND-CIRC straight-line program of at most \( O(s) \) lines.

By \( O(s) \) we mean that the bound is at most \( c \cdot s \) where \( c \) is a constant that is independent of \( n \). For example, if \( f \) can be computed by a Boolean circuit of \( s \) gates, then it can be computed by a NAND-CIRC program of at most \( 3s \) lines, and if \( f \) can be computed by a NAND
circuit of $s$ gates, then it can be computed by an AON-CIRC program of at most $2s$ lines.

**Proof Idea:**

We omit the formal proof, which is obtained by combining Theorem 3.9, Theorem 3.12, and Theorem 3.17. The key observation is that the results we have seen allow us to translate a program/circuit that computes $f$ in one of the above models into a program/circuit that computes $f$ in another model by increasing the lines/gates by at most a constant factor (in fact this constant factor is at most 3).

Theorem 3.9 is a special case of a more general result. We can consider even more general models of computation, where instead of AND/OR/NOT or NAND, we use other operations (see Section 3.6.1 below). It turns out that Boolean circuits are equivalent in power to such models as well. The fact that all these different ways to define computation lead to equivalent models shows that we are “on the right track”. It justifies the seemingly arbitrary choices that we’ve made of using AND/OR/NOT or NAND as our basic operations, since these choices do not affect the power of our computational model. Equivalence results such as Theorem 3.19 mean that we can easily translate between Boolean circuits, NAND circuits, NAND-CIRC programs and the like. We will use this ability later on in this book, often shifting to the most convenient formulation without making a big deal about it. Hence we will not worry too much about the distinction between, for example, Boolean circuits and NAND-CIRC programs.

In contrast, we will continue to take special care to distinguish between circuits/programs and functions (recall Big Idea 2). A function corresponds to a specification of a computational task, and it is a fundamentally different object than a program or a circuit, which corresponds to the implementation of the task.

### 3.6.1 Circuits with other gate sets

There is nothing special about AND/OR/NOT or NAND. For every set of functions $\mathcal{G} = \{G_0, \ldots, G_{k-1}\}$, we can define a notion of circuits that use elements of $\mathcal{G}$ as gates, and a notion of a “$\mathcal{G}$ programming language” where every line involves assigning to a variable foo the result of applying some $G_i \in \mathcal{G}$ to previously defined or input variables. Specifically, we can make the following definition:

### Definition 3.20 — General straight-line programs.

Let $\mathcal{F} = \{f_0, \ldots, f_{t-1}\}$ be a finite collection of Boolean functions, such that $f_i : \{0, 1\}^{k_i} \to \{0, 1\}$ for some $k_i \in \mathbb{N}$. An $\mathcal{F}$ program is a sequence of lines, each of
which assigns to some variable the result of applying some \( f_i \in \mathcal{F} \) to \( k_i \) other variables. As above, we use \( X[i] \) and \( Y[j] \) to denote the input and output variables.

We say that \( \mathcal{F} \) is a universal set of operations (also known as a universal gate set) if there exists a \( \mathcal{F} \) program to compute the function NAND.

AON-CIRC programs correspond to \( \{AND, OR, NOT\} \) programs, NAND-CIRC programs corresponds to \( \mathcal{F} \) programs for the set \( \mathcal{F} \) that only contains the NAND function, but we can also define \( \{IF, ZERO, ONE\} \) programs (see below), or use any other set.

We can also define \( \mathcal{F} \) circuits, which will be directed graphs in which each gate corresponds to applying a function \( f_i \in \mathcal{F} \), and will each have \( k_i \) incoming wires and a single outgoing wire. (If the function \( f_i \) is not symmetric, in the sense that the order of its input matters then we need to label each wire entering a gate as to which parameter of the function it corresponds to.) As in Theorem 3.9, we can show that \( \mathcal{F} \) circuits and \( \mathcal{F} \) programs are equivalent. We have seen that for \( \mathcal{F} = \{AND, OR, NOT\} \), the resulting circuits/programs are equivalent in power to the NAND-CIRC programming language, as we can compute NAND using AND/OR/NOT and vice versa. This turns out to be a special case of a general phenomena—the universality of NAND and other gate sets—that we will explore more in depth later in this book.

Example 3.21 — IF,ZERO,ONE circuits. Let \( \mathcal{F} = \{IF, ZERO, ONE\} \)

where \( ZERO : \{0, 1\} \to \{0\} \) and \( ONE : \{0, 1\} \to \{1\} \) are the constant zero and one functions, \(^3 \) and \( IF : \{0, 1\}^3 \to \{0, 1\} \) is the function that on input \( (a, b, c) \) outputs \( b \) if \( a = 1 \) and \( c \) otherwise.

Then \( \mathcal{F} \) is universal.

Indeed, we can demonstrate that \( \{IF, ZERO, ONE\} \) is universal using the following formula for NAND:

\[
NAND(a, b) = IF(a, IF(b, ZERO, ONE), ONE).
\]

There are also some sets \( \mathcal{F} \) that are more restricted in power. For example it can be shown that if we use only AND or OR gates (without NOT) then we do not get an equivalent model of computation. The exercises cover several examples of universal and non-universal gate sets.

3.6.2 Specification vs. implementation (again)

As we discussed in Section 2.5.1, one of the most important distinctions in this book is that of specification versus implementation or separating “what” from “how” (see Fig. 3.31). A function corresponds to the specification of a computational task, that is what output should
be produced for every particular input. A *program* (or circuit, or any other way to specify *algorithms*) corresponds to the *implementation* of *how* to compute the desired output from the input. That is, a program is a set of instructions how to compute the output from the input. Even within the same computational model there can be many different ways to compute the same function. For example, there is more than one NAND-CIRC program that computes the majority function, more than one Boolean circuit to compute the addition function, and so on and so forth.

Confusing specification and implementation (or equivalently *functions* and *programs*) is a common mistake, and one that is unfortunately encouraged by the common programming-language terminology of referring to parts of programs as “functions”. However, in both the theory and practice of computer science, it is important to maintain this distinction, and it is particularly important for us in this book.

**Lecture Recap**

- An *algorithm* is a recipe for performing a computation as a sequence of “elementary” or “simple” operations.
- One candidate definition for “elementary” operations is the set AND, OR and NOT.
- Another candidate definition for an “elementary” operation is the NAND operation. It is an operation that is easily implementable in the physical world in a variety of methods including by electronic transistors.
- We can use NAND to compute many other functions, including majority, increment, and others.
• There are other equivalent choices, including the sets \( \{ AND, OR, NOT \} \) and \( \{ IF, ZERO, ONE \} \).

• We can formally define the notion of a function \( F : \{0, 1\}^n \rightarrow \{0, 1\}^m \) being computable using the NAND-CIRC Programming language.

• For every set of basic operations, the notions of being computable by a circuit and being computable by a straight-line program are equivalent.

3.7 EXERCISES

Exercise 3.1 — Compare 4 bit numbers. Give a Boolean circuit (with AND/OR/NOT gates) that computes the function \( CMP_8 : \{0, 1\}^8 \rightarrow \{0, 1\} \) such that \( CMP_8(a_0, a_1, a_2, a_3, b_0, b_1, b_2, b_3) = 1 \) if and only if the number represented by \( a_0a_1a_2a_3 \) is larger than the number represented by \( b_0b_1b_2b_3 \).

Exercise 3.2 — Compare \( n \) bit numbers. Prove that there exists a constant \( c \) such that for every \( n \) there is a Boolean circuit (with AND/OR/NOT gates) \( C \) of at most \( c \cdot n \) gates that computes the function \( CMP_{2n} : \{0, 1\}^{2n} \rightarrow \{0, 1\} \) such that \( CMP_{2n}(a_0 \cdots a_{n-1}b_0 \cdots b_{n-1}) = 1 \) if and only if the number represented by \( a_0 \cdots a_{n-1} \) is larger than the number represented by \( b_0 \cdots b_{n-1} \).

Exercise 3.3 — OR, NOT is universal. Prove that the set \( \{ OR, NOT \} \) is universal, in the sense that one can compute NAND using these gates.

Exercise 3.4 — AND, OR is not universal. Prove that for every \( n \)-bit input circuit \( C \) that contains only AND, and OR gates, as well as gates that compute the constant functions 0 and 1, \( C \) is monotone, in the sense that if \( x, x' \in \{0, 1\}^n, x_i \leq x'_i \) for every \( i \in [n] \), then \( C(x) \leq C(x') \).

Conclude that the set \( \{ AND, OR, 0, 1 \} \) is not universal.

Exercise 3.5 — XOR is not universal. Prove that for every \( n \)-bit input circuit \( C \) that contains only XOR, gates, as well as gates that compute the constant functions 0 and 1, \( C \) is affine or linear modulo two, in the sense that there exists some \( a \in \{0, 1\}^n \) and \( b \in \{0, 1\} \) such that for every \( x \in \{0, 1\}^n \), \( C(x) = \sum_{i=0}^{n-1} a_ix_i + b \mod 2 \).

Conclude that the set \( \{ XOR, 0, 1 \} \) is not universal.

Exercise 3.6 — MAJ, NOT, 1 is universal. Let \( MAJ : \{0, 1\}^3 \rightarrow \{0, 1\} \) be the majority function. Prove that \( \{ MAJ, NOT, 1 \} \) is a universal set of gates.
Exercise 3.7 — MAJ, NOT is not universal. Prove that \( \{ \text{MAJ}, \text{NOT} \} \) is not a universal set. See footnote for hint.\(^4\)

Exercise 3.8 — NOR is universal. Let NOR : \( \{0, 1\}^2 \rightarrow \{0, 1\} \) defined as \( \text{NOR}(a, b) = \text{NOT}(\text{OR}(a, b)) \). Prove that \( \{ \text{NOR} \} \) is a universal set of gates.

Exercise 3.9 — Lookup is universal. Prove that \( \{ \text{LOOKUP}_1, 0, 1 \} \) is a universal set of gates where 0 and 1 are the constant functions \( \text{LOOKUP}_1 : \{0, 1\}^3 \rightarrow \{0, 1\} \satisfies \text{LOOKUP}_1(a, b, c) \) equals \( a \) if \( c = 0 \) and equals \( b \) if \( c = 1 \).

Exercise 3.10 — Bound on universal basis size (challenge). Prove that for every subset \( B \) of the functions from \( \{0, 1\}^k \) to \( \{0, 1\} \), if \( B \) is universal then there is a \( B \)-circuit of at most \( O(k) \) gates to compute the NAND function (you can start by showing that there is a \( B \) circuit of at most \( O(k^{16}) \) gates).\(^5\)

Exercise 3.11 — Size and inputs / outputs. Prove that for every NAND circuit of size \( s \) with \( n \) inputs and \( m \) outputs, \( s \geq \min\{n/2, m\} \). See footnote for hint.\(^6\)

Exercise 3.12 — Threshold using NANDs. Prove that there is some constant \( c \) such that for every \( n > 1 \), and integers \( a_0, \ldots, a_{n-1}, b \in \{-2^n, -2^n + 1, \ldots, -1, 0, +1, \ldots, 2^n\} \), there is a NAND circuit with at most \( cn^4 \) gates that computes the threshold function \( f_{a_0, \ldots, a_{n-1}, b} : \{0, 1\}^n \rightarrow \{0, 1\} \) that on input \( x \in \{0, 1\}^n \) outputs 1 if and only if \( \sum_{i=0}^{n-1} a_i x_i > b \).

Exercise 3.13 — NANDs from activation functions. We say that a function \( f : \mathbb{R}^2 \rightarrow \mathbb{R} \) is a NAND approximator if it has the following property: for every \( a, b \in \mathbb{R} \), if \( \min\{|a|, |1-a|\} \leq 1/3 \) and \( \min\{|b|, |1-b|\} \leq 1/3 \) then \( |f(a, b) - \text{NAND}(|a|, |b|)| \leq 1/3 \) where we denote by \( |x| \) the integer closest to \( x \). That is, if \( a, b \) are within a distance \( 1/3 \) to \( \{0, 1\} \) then we want \( f(a, b) \) to equal the NAND of the values in \( \{0, 1\} \) that are closest to \( a \) and \( b \) respectively. Otherwise, we do not care what the output of \( f \) is on \( a \) and \( b \).

In this exercise you will show that you can construct a NAND approximator from many common activation functions used in deep neural networks. As a corollary you will obtain that deep neural networks can simulate NAND circuits. Since NAND circuits can also simulate deep neural networks, these two computational models are equivalent to one another.

\(^4\) Hint: Use the fact that \( \text{MAJ}([\pi, \pi, \pi]) = \text{MAJ}(a, b, c) \) to prove that every \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) computable by a circuit with only \( \text{MAJ} \) and \( \text{NOT} \) gates satisfies \( f(0, 0, \ldots, 0) \neq f(1, 1, \ldots, 1) \). Thanks to Nathan Brunelle and David Evans for suggesting this exercise.

\(^5\) Thanks to Alec Sun for solving this problem.

\(^6\) Hint: Use the conditions of Definition 3.5 stipulating that every input vertex has at least one out-neighbor and there are exactly \( m \) output gates. See also Remark 3.7.
1. Show that there is a NAND approximator \( f \) defined as \( f(a, b) = L(ReLU(L'(a, b))) \) where \( L' : \mathbb{R}^2 \to \mathbb{R} \) is an affine function (of the form \( L'(a, b) = \alpha a + \beta b + \gamma \) for some \( \alpha, \beta, \gamma \in \mathbb{R} \)), \( L \) is an affine function (of the form \( L(y) = \alpha y + \beta \) for \( \alpha, \beta \in \mathbb{R} \)), and \( ReLU : \mathbb{R} \to \mathbb{R} \), is the function defined as \( ReLU(x) = \max\{0, x\} \).

2. Show that there is a NAND approximator \( f \) defined as \( f(a, b) = L(\text{sigmoid}(L'(a, b))) \) where \( L', L \) are affine as above and \( \text{sigmoid} : \mathbb{R} \to \mathbb{R} \) is the function defined as \( \text{sigmoid}(x) = e^x/(e^x + 1) \).

3. Show that there is a NAND approximator \( f \) defined as \( f(a, b) = L(\tanh(L'(a, b))) \) where \( L', L \) are affine as above and \( \tanh : \mathbb{R} \to \mathbb{R} \) is the function defined as \( \tanh(x) = (e^x - e^{-x})/(e^x + e^{-x}) \).

4. Prove that for every NAND-circuit \( C \) with \( n \) inputs and one output that computes a function \( g : \{0, 1\}^n \to \{0, 1\} \), if we replace every gate of \( C \) with a NAND-approximator and then invoke the resulting circuit on some \( x \in \{0, 1\}^n \), the output will be a number \( y \) such that \( |y - g(x)| \leq 1/3 \).

Exercise 3.14 — Majority with NANDs efficiently. Prove that there is some constant \( c \) such that for every \( n > 1 \), there is a NAND circuit of at most \( c \cdot n \) gates that computes the function \( \text{MAJ}_n : \{0, 1\}^n \to \{0, 1\} \) is the majority function on \( n \) input bits. That is \( \text{MAJ}_n(x) = 1 \iff \sum_{i=0}^{n-1} x_i > n/2 \). See footnote for hint.\(^7\)

Exercise 3.15 — Output at last layer. Prove that for every \( f : \{0, 1\}^n \to \{0, 1\} \), if there is a Boolean circuit \( C \) of \( s \) gates that computes \( f \) then there is a Boolean circuit \( C' \) of at most \( s \) gates such that in the minimal layering of \( C' \), the output gate of \( C' \) is in placed the last layer. See footnote for hint.\(^8\)

3.8 BIOGRAPHICAL NOTES


Charles Babbage (1791-1871) was a visionary scientist, mathematician, and inventor (see [Swa02; CM00]). More than a century before the invention of modern electronic computers, Babbage realized that computation can be in principle mechanized. His first design for a mechanical computer was the difference engine that was designed to do polynomial interpolation. He then designed the analytical engine which was a much more general machine and the first prototype for a programmable general purpose computer. Unfortunately, Babbage was

\(^7\) One approach to solve this is using recursion and analyzing it using the so called “Master Theorem”.

\(^8\) Hint: Vertices in layers beyond the output can be safely removed without changing the functionality of the circuit.
never able to complete the design of his prototypes. One of the earliest people to realize the engine’s potential and far reaching implications was Ada Lovelace (see the notes for Chapter 6).

Boolean algebra was first investigated by Boole and DeMorgan in the 1840’s [Boo47; De 47]. The definition of Boolean circuits and connection to electrical relay circuits was given in Shannon’s Masters Thesis [Sha38]. (Howard Gardener called Shannon’s thesis “possibly the most important, and also the most famous, master’s thesis of the [20th] century”.) Savage’s book [Sav98], like this one, introduces the theory of computation starting with Boolean circuits as the first model. Jukna’s book [Juk12] contains a modern in-depth exposition of Boolean circuits, see also [Weg87].

The NAND function was shown to be universal by Sheffer [She13], though this also appears in the earlier work of Peirce, see [Bur78]. Whitehead and Russell used NAND as the basis for their logic in their magnum opus Principia Mathematica [WR12]. In her Ph.D thesis, Ernst [Ern09] investigates empirically the minimal NAND circuits for various functions. Nissan and Shocken’s book [NS05] builds a computing system starting from NAND gates and ending with high level programs and games (“NAND to Tetris”); see also the website nandtotetris.org.
Syntactic sugar, and computing every function

“The [In 1951] I had a running compiler and nobody would touch it because, they carefully told me, computers could only do arithmetic; they could not do programs.”, Grace Murray Hopper, 1986.

“Syntactic sugar causes cancer of the semicolon.”, Alan Perlis, 1982.

The computational models we considered thus far are as “bare bones” as they come. For example, our NAND-CIRC “programming language” has only the single operation $\text{foo} = \text{NAND}(\text{bar}, \text{blah})$. In this chapter we will see that these simple models are actually equivalent to more powerful ones. The key observation is that we can implement more complex features using our basic building blocks, and then use these new features themselves as building blocks for even more sophisticated features. This is known as “syntactic sugar” in the field of programming language design, since we are not modifying the underlying programming model itself but rather we merely implement new features by syntactically transforming a program that uses such features into one that doesn’t.

This chapter provides a “toolkit” that can be used to show that many functions can be computed by NAND-CIRC programs, and hence also by Boolean circuits. We will also use this toolkit to prove a fundamental theorem: every finite function $f : \{0, 1\}^n \rightarrow \{0, 1\}^m$ can be computed by a Boolean circuit, see Theorem 4.13 below. While the syntactic sugar toolkit is important in its own right, Theorem 4.13 can also be proven directly without using this toolkit. We present this alternative proof in Section 4.5. See Fig. 4.1 for an outline of the results of this chapter.
We now present some examples of “syntactic sugar” transformations that we can use in constructing straightline programs or circuits. We focus on the straight-line programming language view of our computational models, and specifically (for the sake of concreteness) on the NAND-CIRC programming language. This is convenient because many of the syntactic sugar transformations we present are easiest to think about in terms of applying “search and replace” operations to the source code of a program. However, by Theorem 3.19, all of our results hold equally well for circuits, whether ones using NAND gates or Boolean circuits that use the AND, OR, and NOT operations. Enumerating the examples of such syntactic sugar transformations can be a little tedious, but we do it for two reasons:

1. To convince you that despite their seeming simplicity and limitations, simple models such as Boolean circuits or the NAND-CIRC programming language are actually quite powerful.

2. So you can realize how lucky you are to be taking a theory of computation course and not a compilers course... :)

4.1 SOME EXAMPLES OF SYNTACTIC SUGAR

4.1.1 User-defined procedures

One staple of almost any programming language is the ability to define and then execute procedures or subroutines. (These are often known as functions in some programming languages, but we prefer the name procedures to avoid confusion with the function that a program computes.) The NAND-CIRC programming language does not have this mechanism built in. However, we can achieve the same
effect using the time honored technique of “copy and paste”. Specifically, we can replace code which defines a procedure such as

```python
def Proc(a, b):
    proc_code
    return c
some_code
f = Proc(d, e)
some_more_code
```

with the following code where we “paste” the code of `Proc`

```python
some_code
proc_code'
some_more_code
```

and where `proc_code'` is obtained by replacing all occurrences of `a` with `d`, `b` with `e`, and `c` with `f`. When doing that we will need to ensure that all other variables appearing in `proc_code'` don’t interfere with other variables. We can always do so by renaming variables to new names that were not used before. The above reasoning leads to the proof of the following theorem:

**Theorem 4.1 — Procedure definition syntactic sugar.** Let NAND-CIRC-PROC be the programming language NAND-CIRC augmented with the syntax above for defining procedures. Then for every NAND-CIRC-PROC program $P$, there exists a standard (i.e., “sugar free”) NAND-CIRC program $P'$ that computes the same function as $P$.

**Remark 4.2 — No recursive procedure.** NAND-CIRC-PROC only allows non recursive procedures. In particular, the code of a procedure `Proc` can not call `Proc` but only use procedures that were defined before it. Without this restriction, the above “search and replace” procedure might never terminate and Theorem 4.1 would not be true.

Theorem 4.1 can be proven using the transformation above, but since the formal proof is somewhat long and tedious, we omit it here.

**Example 4.3 — Computing Majority from NAND using syntactic sugar.** Procedures allow us to express NAND-CIRC programs much more cleanly and succinctly. For example, because we can compute
AND, OR, and NOT using NANDs, we can compute the Majority function as follows:

```python
def NOT(a):
    return NAND(a,a)
def AND(a,b):
    temp = NAND(a,b)
    return NOT(temp)
def OR(a,b):
    temp1 = NOT(a)
    temp2 = NOT(b)
    return NAND(temp1,temp2)

def MAJ(a,b,c):
    and1 = AND(a,b)
    and2 = AND(a,c)
    and3 = AND(b,c)
    or1 = OR(and1,and2)
    return OR(or1,and3)

print(MAJ(0,1,1))  # 1
```

Fig. 4.2 presents the “sugar free” NAND-CIRC program (and the corresponding circuit) that is obtained by “expanding out” this program, replacing the calls to procedures with their definitions.

**Big Idea 4** Once we show that a computational model $X$ is equivalent in power to the model with an additional feature $Y$, we can use this feature whenever we need to show that some function $f$ is computable by $X$.

**Remark 4.4 — Counting lines.** While we can use syntactic sugar to present NAND-CIRC programs in more readable ways, we did not change the definition of the language itself. Therefore, whenever we say that some function $f$ has an $s$-line NAND-CIRC program we mean a standard “sugar free” NAND-CIRC program, where all syntactic sugar has been expanded out. For example, the program of Example 4.3 is a 12-line program for computing the MAJ function, even though it can be written in fewer lines using NAND-CIRC-PROC.
4.1.2 Proof by Python (optional)

We can write a Python program that implements the proof of Theorem 4.1. This is a Python program that takes a NAND-CIRC-PROC program $P$ that includes procedure definitions and uses simple “search and replace” to transform $P$ into a standard (i.e., “sugar free”) NAND-CIRC program $P'$ that computes the same function as $P$ without using any procedures. The idea is simple: if the program $P$ contains a definition of a procedure $Proc$ of two arguments $x$ and $y$, then whenever we see a line of the form $foo = Proc(bar,blah)$, we can replace this line by:

1. The body of the procedure $Proc$ (replacing all occurrences of $x$ and $y$ with $bar$ and $blah$ respectively).

2. A line $foo = exp$, where $exp$ is the expression following the return statement in the definition of the procedure $Proc$.

To make this more robust we a prefix to the internal variables used by $Proc$ to ensure they don’t conflict with the variables of $P$; for simplicity we ignore this issue in the code below though it can be easily added.

The code in ?? achieves such a transformation.\(^1\)

```python
def inline_proc(code, proc_name, proc_args, proc_body):
    '''Takes code of a program and name, arguments, body of a procedure.
    Returns new code where all lines in program of the form "foo = proc_name(bar,blah...)" are replaced with the body of the procedure with arguments instantiated with the variables bar, blah, etc.''

    arglist = ",".join([r"([a-zA-Z0-9\._\:\[\]\:\-]+)" for i in range(len(proc_args))])
```

---

\(^1\) This code uses regular expressions to make the search and replace parts a little easier. We will see the theoretical basis for regular expressions in Chapter 9.
Fig. 4.2 shows the result of applying the code of ?? to the program of Example 4.3 that uses syntactic sugar to compute the Majority function. Specifically, we first apply desugar to remove usage of the OR function, then apply it to remove usage of the AND function, and finally apply it a third time to remove usage of the NOT function.

Remark 4.5 — Parsing function definitions (optional). The function desugar in ?? assumes that it is given the procedure already split up into its name, arguments, and body. It is not crucial for our purposes to describe precisely to scan a definition and splitting it up to these components, but in case you are curious, it can be achieved in Python via the following code:

```python
def parse_procs(code):
    lines = [l for l in code.split('\n') if l]
    regexp = r'def\s+([a-zA-Z0-9-_:]+)\(([^,]+)\)\s*:\s*'
    procs = []
    current_line = 0
    rest = ''
    while current_line < len(lines):
        m = re.match(regexp, lines[current_line])
        if m:
            current_line += 1
            code = ''
            while current_line < len(lines) and lines[current_line][0] == ':':
                code += lines[current_line]
                current_line += 1
            procs.append((m.group(1), m.group(2), code))
    return procs
```

### Syntax Highlighting

```python
def parse_procs(code):
    lines = [l for l in code.split('n') if l]
    regexp = r'def\s+([a-zA-Z0-9-_:]+)\(([^,]+)\)\s*:\s*'
    procs = []
    current_line = 0
    rest = ''
    while current_line < len(lines):
        m = re.match(regexp, lines[current_line])
        if m:
            current_line += 1
            code = ''
            while current_line < len(lines) and lines[current_line][0] == ':':
                code += lines[current_line]
                current_line += 1
            procs.append((m.group(1), m.group(2), code))
    return procs
```
4.1.3 Conditional statements

Another sorely missing feature in NAND-CIRC is a conditional statement such as the if/then constructs that are found in many programming languages. However, using procedures, we can obtain an ersatz if/then construct. First we can compute the function $IF: \{0,1\}^3 \rightarrow \{0,1\}$ such that $IF(a, b, c)$ equals $b$ if $a = 1$ and $c$ if $a = 0$.

Before reading onward, try to see how you could compute the $IF$ function using NAND’s. Once you do that, see how you can use that to emulate if/then types of constructs.

The $IF$ function can be implemented from NANDs as follows (see Exercise 4.2):

```python
def IF(cond, a, b):
    notcond = NAND(cond, cond)
    temp = NAND(b, notcond)
    temp1 = NAND(a, cond)
    return NAND(temp, temp1)
```

The $IF$ function is also known as a multiplexing function, since $cond$ can be thought of as a switch that controls whether the output is connected to $a$ or $b$. Once we have a procedure for computing the $IF$ function, we can implement conditionals in NAND. The idea is that we replace code of the form

```python
if (condition): assign blah to variable foo
```

with code of the form

```python
foo = IF(condition, blah, foo)
```

that assigns to $foo$ its old value when condition equals 0, and assign to $foo$ the value of $blah$ otherwise. More generally we can replace code of the form...
if (cond):
    a = ...
    b = ...
    c = ...

with code of the form

temp_a = ...
temp_b = ...
temp_c = ...
a = IF(cond, temp_a, a)
b = IF(cond, temp_b, b)
c = IF(cond, temp_c, c)

Using such transformations, we can prove the following theorem. Once again we omit the (not too insightful) full formal proof, though see Section 4.1.2 for some hints on how to obtain it.

**Theorem 4.6 — Conditional statements syntactic sugar.** Let NAND-CIRC-IF be the programming language NAND-CIRC augmented with if/then/else statements for allowing code to be conditionally executed based on whether a variable is equal to 0 or 1.

Then for every NAND-CIRC-IF program $P$, there exists a standard (i.e., “sugar free”) NAND-CIRC program $P'$ that computes the same function as $P$.

### 4.2 Extended Example: Addition and Multiplication (Optional)

Using “syntactic sugar”, we can write the integer addition function as follows:

```python
# Add two n-bit integers  
# Use LSB first notation for simplicity

def ADD(A,B):
    Result = [0]*(n+1)
    Carry = [0]*(n+1)
    Carry[0] = zero(A[0])
    for i in range(n):
        Result[i] = XOR(Carry[i], XOR(A[i], B[i]))
        Carry[i+1] = MAJ(Carry[i], A[i], B[i])
    Result[n] = Carry[n]
    return Result

ADD([1,1,1,0,0],[1,0,0,0,0]);
# [0, 0, 0, 1, 0, 0]
```
where zero is the constant zero function, and MAJ and XOR correspond to the majority and XOR functions respectively. In the above we used the loop for i in range(n) but we can expand this out by simply repeating the code n times, replacing the value of i with 0, 1, 2, …, n − 1. By expanding out all the features, for every value of n we can translate the above program into a standard (“sugar free”) NAND-CIRC program. Fig. 4.3 depicts what we get for n = 2.

By going through the above program carefully and accounting for the number of gates, we can see that it yields a proof of the following theorem (see also Fig. 4.4):

**Theorem 4.7 — Addition using NAND-CIRC programs.** For every \(n \in \mathbb{N}\), let \(ADD_n : \{0, 1\}^{2n} \to \{0, 1\}^{n+1}\) be the function that, given \(x, x' \in \{0, 1\}^n\) computes the representation of the sum of the numbers that \(x\) and \(x'\) represent. Then there is a constant \(c \leq 30\) such that for every \(n\) there is a NAND-CIRC program of at most \(cn\) lines computing \(ADD_n\).

Once we have addition, we can use the grade-school algorithm to obtain multiplication as well, thus obtaining the following theorem:

**Theorem 4.8 — Multiplication using NAND-CIRC programs.** For every \(n\), let \(MULT_n : \{0, 1\}^{2n} \to \{0, 1\}^{2n}\) be the function that, given \(x, x' \in \{0, 1\}^n\) computes the representation of the product of the numbers that \(x\) and \(x'\) represent. Then there is a constant \(c\) such that for every \(n\), there is a NAND-CIRC program of at most \(cn^2\) that computes the function \(MULT_n\).
We omit the proof, though in Exercise 4.7 we ask you to supply a “constructive proof” in the form of a program (in your favorite programming language) that on input a number $n$, outputs the code of a NAND-CIRC program of at most $1000n^2$ lines that computes the $MULT_n$ function. In fact, we can use Karatsuba’s algorithm to show that there is a NAND-CIRC program of $O(n \log_2^3)$ lines to compute $MULT_n$ (and can get even further asymptotic improvements using better algorithms).

### 4.3 THE LOOKUP FUNCTION

The $LOOKUP$ function will play an important role in this chapter and later. It is defined as follows:

**Definition 4.9 — Lookup function.** For every $k$, the $lookup$ function of order $k$, $LOOKUP_k : \{0, 1\}^{2^k + k} \rightarrow \{0, 1\}$ is defined as follows. For every $x \in \{0, 1\}^{2^k}$ and $i \in \{0, 1\}^k$,

$$LOOKUP_k(x, i) = x_i$$  \hspace{1cm} (4.1)

where $x_i$ denotes the $i^{th}$ entry of $x$, using the binary representation to identify $i$ with a number in $\{0, \ldots, 2^k - 1\}$.

See Fig. 4.5 for an illustration of the $LOOKUP$ function. It turns out that for every $k$, we can compute $LOOKUP_k$ using a NAND-CIRC program:

**Theorem 4.10 — Lookup function.** For every $k > 0$, there is a NAND-CIRC program that computes the function $LOOKUP_k : \{0, 1\}^{2^k + k} \rightarrow \{0, 1\}$. Moreover, the number of lines in this program is at most $4 \cdot 2^k$.

An immediate corollary of Theorem 4.10 is that for every $k > 0$, $LOOKUP_k$ can be computed by a Boolean circuit (with AND, OR and NOT gates) of at most $8 \cdot 2^k$ gates.
4.3.1 Constructing a NAND-CIRC program for \textit{LOOKUP}

We prove Theorem 4.10 by induction. For the case $k = 1$, LOOKUP\textsubscript{1} maps $(x_0, x_1, i) \in \{0, 1\}^3$ to $x_i$. In other words, if $i = 0$ then it outputs $x_0$ and otherwise it outputs $x_1$, which (up to reordering variables) is the same as the IF function presented in Section 4.1.3, which can be computed by a 4-line NAND-CIRC program.

As a warm-up for the case of general $k$, let us consider the case of $k = 2$. Given input $x = (x_0, x_1, x_2, x_3)$ for \textit{LOOKUP}\textsubscript{2} and an index $i = (i_0, i_1)$, if the most significant bit $i_0$ of the index is 0 then \textit{LOOKUP}\textsubscript{2}(x, i) will equal $x_0$ if $i_1 = 0$ and equal $x_1$ if $i_1 = 1$. Similarly, if the most significant bit $i_0$ is 1 then \textit{LOOKUP}\textsubscript{2}(x, i) will equal $x_2$ if $i_1 = 0$ and will equal $x_3$ if $i_1 = 1$. Another way to say this is that we can write \textit{LOOKUP}\textsubscript{2} as follows:

\begin{verbatim}
def LOOKUP2(X[0], X[1], X[2], X[3], i[0], i[1]):
    if i[0]==1:
        return LOOKUP1(X[2], X[3], i[1])
    else:
        return LOOKUP1(X[0], X[1], i[1])
\end{verbatim}

or in other words,

\begin{verbatim}
def LOOKUP2(X[0], X[1], X[2], X[3], i[0], i[1]):
    a = LOOKUP1(X[2], X[3], i[1])
    b = LOOKUP1(X[0], X[1], i[1])
    return IF( i[0], a, b)
\end{verbatim}

More generally, as shown in the following lemma, we can compute \textit{LOOKUP}\textsubscript{k} using two invocations of \textit{LOOKUP}\textsubscript{k−1} and one invocation of IF:

\textbf{Lemma 4.11 — Lookup recursion.} For every $k \geq 2$, \textit{LOOKUP}\textsubscript{k}(x_0, \ldots, x_{2^{k−1}−1}, i_0, \ldots, i_{k−1})

is equal to

\begin{equation}
\text{IF}(i_0, \text{LOOKUP}\textsubscript{k−1}(x_{2^{k−1}}, \ldots, x_{2^{k−1}−1}, i_1, \ldots, i_{k−1}), \text{LOOKUP}\textsubscript{k−1}(x_0, \ldots, x_{2^{k−1}−1}, i_1, \ldots, i_{k−1}))
\end{equation}

(4.2)

\textbf{Proof.} If the most significant bit $i_0$ of $i$ is zero, then the index $i$ is in $\{0, \ldots, 2^{k−1}−1\}$ and hence we can perform the lookup on the “first half” of $x$ and the result of \textit{LOOKUP}\textsubscript{k}(x, i) will be the same as $a = \text{LOOKUP}\textsubscript{k−1}(x_0, \ldots, x_{2^{k−1}−1}, i_1, \ldots, i_{k−1})$. On the other hand, if this most significant bit $i_0$ is equal to 1, then the index is in $\{2^{k−1}, \ldots, 2^k − 1\}$, in which case the result of \textit{LOOKUP}\textsubscript{k}(x, i) is the same as $b = \text{LOOKUP}\textsubscript{k−1}(x_{2^{k−1}}, \ldots, x_{2^{k−1}−1}, i_1, \ldots, i_{k−1})$. Thus we can compute \textit{LOOKUP}\textsubscript{k}(x, i) by first computing $a$ and $b$ and then outputting \textit{IF}(i, a, b).

\blacksquare
Proof of Theorem 4.10 from Lemma 4.11. Now that we have Lemma 4.11, we can complete the proof of Theorem 4.10. We will prove by induction on \( k \) that there is a NAND-CIRC program of at most \( 4 \cdot 2^k \) lines for \( \text{LOOKUP}_k \). For \( k = 1 \) this follows by the four line program for \( \text{IF} \) we’ve seen before. For \( k > 1 \), we use the following pseudocode

\[
\begin{align*}
    a &= \text{LOOKUP}_{k-1}(X[0], \ldots, X[2^{(k-1)}-1], i[0], \ldots, i[k-1]) \\
    b &= \text{LOOKUP}_{k-1}(X[2^{(k-1)}], \ldots, Z[2^{(k-1)}], i[0], \ldots, i[k-1]) \\
    \text{return } &\text{IF}(i[0], b, a)
\end{align*}
\]

If we let \( L(k) \) be the number of lines required for \( \text{LOOKUP}_k \), then the above pseudo-code shows that

\[
L(k) \leq 2L(k-1) + 4 .
\] (4.3)

which solves for \( L(k) \leq 4(2^k - 1) \). See Fig. 4.6 for a plot of the actual number of lines in our implementation of \( \text{LOOKUP}_k \).

4.4 COMPUTING EVERY FUNCTION

At this point we know the following facts about NAND-CIRC programs (and so equivalently about Boolean circuits and our other equivalent models):

1. They can compute at least some non trivial functions.
2. Coming up with NAND-CIRC programs for various functions is a very tedious task.

Thus I would not blame the reader if they were not particularly looking forward to a long sequence of examples of functions that can be computed by NAND-CIRC programs. However, it turns out we are not going to need this, as we can show in one fell swoop that NAND-CIRC programs can compute every finite function:

**Theorem 4.12 — Universality of NAND.** There exists some constant \( c > 0 \) such that for every \( n, m > 0 \) and function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \), there is a NAND-CIRC program with at most \( c \cdot m2^n \) lines that computes the function \( f \).

By Theorem 3.19, the models of NAND circuits, NAND-CIRC programs, AON-CIRC programs, and Boolean circuits, are all equivalent to one another, and hence Theorem 4.12 holds for all these models. In particular, the following theorem is equivalent to Theorem 4.12:
**Theorem 4.13 — Universality of Boolean circuits.** There exists some constant $c > 0$ such that for every $n, m > 0$ and function $f : \{0, 1\}^n \to \{0, 1\}^m$, there is a Boolean circuit with at most $c \cdot m 2^n$ gates that computes the function $f$.

**Big Idea 5** Every finite function can be computed by a large enough Boolean circuit.

*Improved bounds.* Though it will not be of great importance to us, it is possible to improve on the proof of Theorem 4.12 and shave an extra factor of $n$, as well as optimize the constant $c$, and so prove that for every $\epsilon > 0$, $m \in \mathbb{N}$ and sufficiently large $n$, if $f : \{0, 1\}^n \to \{0, 1\}^m$ then $f$ can be computed by a NAND circuit of at most $(1 + \epsilon) \frac{m 2^n}{n}$ gates. The proof of this result is beyond the scope of this book, but we do discuss how to obtain a bound of the form $O(\frac{m 2^n}{n})$ in Section 4.4.2; see also the biographical notes.

### 4.4.1 Proof of NAND’s Universality

To prove Theorem 4.12, we need to give a NAND circuit, or equivalently a NAND-CIRC program, for every possible function. We will restrict our attention to the case of Boolean functions (i.e., $m = 1$). Exercise 4.9 asks you to extend the proof for all values of $m$. A function $F : \{0, 1\}^n \to \{0, 1\}$ can be specified by a table of its values for each one of the $2^n$ inputs. For example, the table below describes one particular function $G : \{0, 1\}^4 \to \{0, 1\}$.

Table 4.1: An example of a function $G : \{0, 1\}^4 \to \{0, 1\}$.\(^3\)

<table>
<thead>
<tr>
<th>Input $(x)$</th>
<th>Output $(G(x))$</th>
</tr>
</thead>
<tbody>
<tr>
<td>0000</td>
<td>1</td>
</tr>
<tr>
<td>0001</td>
<td>1</td>
</tr>
<tr>
<td>0010</td>
<td>0</td>
</tr>
<tr>
<td>0011</td>
<td>0</td>
</tr>
<tr>
<td>0100</td>
<td>1</td>
</tr>
<tr>
<td>0101</td>
<td>0</td>
</tr>
<tr>
<td>0110</td>
<td>0</td>
</tr>
<tr>
<td>0111</td>
<td>1</td>
</tr>
<tr>
<td>1000</td>
<td>0</td>
</tr>
<tr>
<td>1001</td>
<td>0</td>
</tr>
<tr>
<td>1010</td>
<td>0</td>
</tr>
<tr>
<td>1011</td>
<td>0</td>
</tr>
</tbody>
</table>

\(^3\)In case you are curious, this is the function on input $i \in \{0, 1\}^4$ (which we interpret as a number in $[16]$), that outputs the $i$-th digit of $\pi$ in the binary basis.
For every $x \in \{0, 1\}^4$, $G(x) = \text{LOOKUP}_4(1100100100001111, x)$, and so the following is NAND-CIRC “pseudocode” to compute $G$ using syntactic sugar for the \text{LOOKUP}_4 procedure.

\[
\begin{array}{c}
G0000 = 1 \\
G1000 = 1 \\
G0100 = 0 \\
\ldots \\
G0111 = 1 \\
G1111 = 1 \\
Y[0] = \text{LOOKUP}_4(G0000, G1000, \ldots, G1111, \\
X[0], X[1], X[2], X[3])
\end{array}
\]

We can translate this pseudocode into an actual NAND-CIRC program by adding three lines to define variables \text{zero} and \text{one} that are initialized to 0 and 1 respectively, and then replacing a statement such as $Gxxx = 0$ with $Gxxx = \text{NAND}($\text{one}, \text{one}$)$ and a statement such as $Gxxx = 1$ with $Gxxx = \text{NAND}($\text{zero}, \text{zero}$)$. The call to \text{LOOKUP}_4 will be replaced by the NAND-CIRC program that computes $LOOKUP_4$, plugging in the appropriate inputs.

There was nothing about the above reasoning that was particular to the function $G$ of \??. Given every function $F : \{0, 1\}^n \rightarrow \{0, 1\}$, we can write a NAND-CIRC program that does the following:

1. Initialize $2^n$ variables of the form $F0\ldots0$ till $F1\ldots1$ so that for every $z \in \{0, 1\}^n$, the variable corresponding to $z$ is assigned the value $F(z)$.

2. Compute $LOOKUP_n$ on the $2^n$ variables initialized in the previous step, with the index variable being the input variables $X[0]$, $X[1]$, $\ldots$, $X[2^n - 1]$. That is, just like in the pseudocode for $G$ above, we use $Y[0] = \text{LOOKUP}(F0\ldots0, \ldots, F1\ldots1, \ldots, X[0], \ldots, X[n-1])$

The total number of lines in the resulting program is $3 + 2^n$ lines for initializing the variables plus the $4 \cdot 2^n$ lines that we pay for computing $LOOKUP_n$. This completes the proof of Theorem 4.12.
s

Remark 4.14 — Result in perspective. While Theorem 4.12 seems striking at first, in retrospect, it is perhaps not that surprising that every finite function can be computed with a NAND-CIRC program. After all, a finite function \( F : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be represented by simply the list of its outputs for each one of the \( 2^n \) input values. So it makes sense that we could write a NAND-CIRC program of similar size to compute it. What is more interesting is that some functions, such as addition and multiplication, have a much more efficient representation: one that only requires \( O(n^2) \) or even fewer lines.

4.4.2 Improving by a factor of \( n \) (optional)

By being a little more careful, we can improve the bound of Theorem 4.12 and show that every function \( F : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be computed by a NAND-CIRC program of at most \( O(m2^n/n) \) lines. In other words, we can prove the following improved version:

Theorem 4.15 — Universality of NAND circuits, improved bound. There exists a constant \( c > 0 \) such that for every \( n, m > 0 \) and function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \), there is a NAND-CIRC program with at most \( c \cdot m2^n/n \) lines that computes the function \( f \).

Proof. As before, it is enough to prove the case that \( m = 1 \). Hence we let \( f : \{0, 1\}^n \rightarrow \{0, 1\} \), and our goal is to prove that there exists a NAND-CIRC program of \( O(2^n/n) \) lines (or equivalently a Boolean circuit of \( O(2^n/n) \) gates) that computes \( f \).

We let \( k = \log(n - 2 \log n) \) (the reasoning behind this choice will become clear later on). We define the function \( g : \{0, 1\}^k \rightarrow \{0, 1\}^{2^{n-k}} \) as follows:

\[
g(a) = f(a0^{n-k})f(a0^{n-k-1}1) \cdots f(a1^{n-k}) . \tag{4.4}
\]

In other words, if we use the usual binary representation to identify the numbers \( \{0, \ldots, 2^{n-k} - 1\} \) with the strings \( \{0, 1\}^{n-k} \), then for every \( a \in \{0, 1\}^k \) and \( b \in \{0, 1\}^{n-k} \)

\[
g(a)_b = f(ab) . \tag{4.5}
\]

(4.6) means that for every \( x \in \{0, 1\}^n \), if we write \( x = ab \) with \( a \in \{0, 1\}^k \) and \( b \in \{0, 1\}^{n-k} \) then we can compute \( f(x) \) by first computing the string \( T = g(a) \) of length \( 2^{n-k} \), and then computing \( \text{LOOKUP}_{n-k}(T, b) \) to retrieve the element of \( T \) at the position corresponding to \( b \) (see Fig. 4.7). The cost to compute the \( \text{LOOKUP}_{n-k} \) is \( O(2^{n-k}) \) lines/gates and the cost in NAND-CIRC lines (or Boolean

![Figure 4.7](image-url)
gates) to compute $f$ is at most

$$
cost(g) + O(2^{n-k}) ,
$$
(4.6)

where $\text{cost}(g)$ is the number of operations (i.e., lines of NAND-CIRC programs or gates in a circuit) needed to compute $g$.

To complete the proof we need to give a bound on $\text{cost}(g)$. Since $g$ is a function mapping $\{0, 1\}^k$ to $\{0, 1\}^{2^{n-k}}$, we can also think of it as a collection of $2^{n-k}$ functions $g_0, \ldots, g_{2^{n-k}-1} : \{0, 1\}^k \to \{0, 1\}$, where $g_i(x) = g(a)_i$ for every $a \in \{0, 1\}^k$ and $i \in [2^{n-k}]$. (That is, $g_i(a)$ is the $i$-th bit of $g(a)$.) Naively, we could use Theorem 4.12 to compute each $g_i$ in $O(2^k)$ lines, but then the total cost is $O(2^{n-k} \cdot 2^k) = O(2^n)$ which does not save us anything. However, the crucial observation is that there are only $2^{2k}$ distinct functions mapping $\{0, 1\}^k$ to $\{0, 1\}$. For example, if $g_{17}$ is an identical function to $g_{67}$ that means that if we already computed $g_{17}(a)$ then we can compute $g_{67}(a)$ using only a constant number of operations: simply copy the same value! In general, if you have a collection of $N$ functions $g_0, \ldots, g_{N-1}$ mapping $\{0, 1\}^k$ to $\{0, 1\}$, of which at most $S$ are distinct then for every value $a \in \{0, 1\}^k$ we can compute the $N$ values $g_0(a), \ldots, g_{N-1}(a)$ using at most $O(S \cdot 2^k + N)$ operations (see Fig. 4.8).

In our case, because there are at most $2^{2k}$ distinct functions mapping $\{0, 1\}^k$ to $\{0, 1\}$, we can compute the function $g$ (and hence by (4.6) also $f$) using at most

$$
O(2^{2k} \cdot 2^k + 2^{n-k})
$$
(4.7)

operations. Now all that is left is to plug into (4.7) our choice of $k = \log(n - 2 \log n)$. By definition, $2^k = n - 2 \log n$, which means that (4.7) can be bounded

$$
O \left( 2^{n-2 \log n} \cdot (n - 2 \log n) + 2^{n-\log(n-2 \log n)} \right) \leq
$$
(4.8)

$$
O \left( \frac{2^n}{n^2} \cdot n + \frac{2^n}{n-2 \log n} \right) \leq O \left( \frac{2^n}{n^2} + \frac{2^n}{n-2 \log n} \right) = O \left( \frac{2^n}{n^2} \right)
$$
(4.9)

which is what we wanted to prove. (We used above the fact that $n - 2 \log n \geq 0.5 \log n$ for sufficiently large $n$.)

Using the connection between NAND-CIRC programs and Boolean circuits, an immediate corollary of Theorem 4.15 is the following improvement to Theorem 4.13:

**Theorem 4.16 — Universality of Boolean circuits, improved bound.** There exists some constant $c > 0$ such that for every $n, m > 0$ and func-

![Figure 4.8](image)
4.5 COMPUTING EVERY FUNCTION: AN ALTERNATIVE PROOF

Theorem 4.13 is a fundamental result in the theory (and practice!) of computation. In this section we present an alternative proof of this basic fact that Boolean circuits can compute every finite function. This alternative proof gives a somewhat worse quantitative bound on the number of gates but it has the advantage of being simpler, working directly with circuits and avoiding the usage of all the syntactic sugar machinery. (However, that machinery is useful in its own right, and will find other applications later on.)

**Theorem 4.17 — Universality of Boolean circuits (alternative phrasing).** There exists some constant $c > 0$ such that for every $n, m > 0$ and function $f : \{0, 1\}^n \to \{0, 1\}^m$, there is a Boolean circuit with at most $c \cdot m \cdot 2^n$ gates that computes the function $f$.

**Proof Idea:**

The idea of the proof is illustrated in Fig. 4.9. As before, it is enough to focus on the case that $m = 1$ (the function $f$ has a single output), since we can always extend this to the case of $m > 1$ by looking at the composition of $m$ circuits each computing a different output bit of the function $f$. We start by showing that for every $\alpha \in \{0, 1\}^n$, there is an $O(n)$ sized circuit that computes the function $\delta_\alpha : \{0, 1\}^n \to \{0, 1\}$ defined as follows: $\delta_\alpha(x) = 1$ iff $x = \alpha$ (that is, $\delta_\alpha$ outputs 0 on all inputs except the input $\alpha$). We can then write any function
f : {0, 1}^n → {0, 1} as the OR of at most 2^n functions δ_α for the α’s on which f(α) = 1.

* Proof of Theorem 4.17. We prove the theorem for the case m = 1. The result can be extended for m > 1 as before (see also Exercise 4.9). Let f : {0, 1}^n → {0, 1}. We will prove that there is an O(n ⋅ 2^n)-sized Boolean circuit to compute f in the following steps:

1. We show that for every α ∈ {0, 1}^n, there is an O(n) sized circuit that computes the function δ_α : {0, 1}^n → {0, 1}, where δ_α(x) = 1 iff x = α.

2. We then show that this implies the existence of an O(n ⋅ 2^n)-sized circuit that computes f, by writing f(x) as the OR of δ_α(x) for all α ∈ {0, 1}^n such that f(α) = 1.

We start with Step 1:
CLAIM: For α ∈ {0, 1}^n, define δ_α : {0, 1}^n as follows:

\[ δ_α(x) = \begin{cases} 
1 & x = α \\
0 & \text{otherwise}
\end{cases} \quad (4.10) \]

then there is a Boolean circuit using at most 2n gates that computes δ_α.

PROOF OF CLAIM: The proof is illustrated in ??.. As an example, consider the function δ_{011} : {0, 1}^3 → {0, 1}. This function outputs 1 on x if and only if x_0 = 0, x_1 = 1 and x_2 = 1, and so we can write δ_{011}(x) = x_0 ∧ x_1 ∧ x_2, which translates into a Boolean circuit with one NOT gate and two AND gates. More generally, for every α ∈ {0, 1}^n, we can express δ_α(x) as (x_0 = α_0) ∧ (x_1 = α_1) ∧ ⋯ ∧ (x_{n−1} = α_{n−1}),

where if α_i = 0 we replace x_i = α_i with x_i and if α_i = 1 we replace x_i = α_i by simply x_i. This yields a circuit that computes δ_α using n AND gates and at most n NOT gates, so a total of at most 2n gates.

Now for every function f : {0, 1}^n → {0, 1}, we can write

\[ f(x) = δ_{x_0}(x) ∨ δ_{x_1}(x) ∨ ⋯ ∨ δ_{x_{n−1}}(x) \quad (4.11) \]

where S = \{x_0, …, x_{n−1}\} is the set of inputs on which f outputs 1.

(To see this, you can verify that the right-hand side of (4.11) evaluates to 1 on x ∈ {0, 1}^n if and only if x is in the set S.)

Therefore we can compute f using a Boolean circuit of at most 2n gates for each of the N functions δ_{x_i} and combine that with at most N OR gates, thus obtaining a circuit of at most 2n ⋅ N + N gates. Since S ⊆ {0, 1}^n, its size N is at most 2^n and hence the total number of gates in this circuit is O(n ⋅ 2^n).


4.6 THE CLASS SIZE(T)

We have seen that every function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be computed by a circuit of size \( O(m \cdot 2^n) \), and some functions (such as addition and multiplication) can be computed by much smaller circuits. We define \( \text{SIZE}(s) \) to be the set of functions that can be computed by NAND circuits of at most \( s \) gates (or equivalently, by NAND-CIRC programs of at most \( s \) lines). Formally, the definition is as follows:

**Definition 4.18 — Size class of functions.** For every \( n, m \in \{1, \ldots, 2s\} \), we let set \( \text{SIZE}_{n,m}(s) \) denotes the set of all functions \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) such that \( f \in \text{SIZE}(s) \). We denote by \( \text{SIZE}_n(s) \) the set \( \text{SIZE}_{n,1}(s) \). For every integer \( s \geq 1 \), we let \( \text{SIZE}(s) = \bigcup_{n,m \leq 2s} \text{SIZE}_{n,m}(s) \) be the set of all functions \( f \) for which there exists a NAND circuit of at most \( s \) gates that compute \( f \).

?? depicts the sets \( \text{SIZE}_{n,1}(s) \). Note that \( \text{SIZE}_{n,m}(s) \) is a set of functions, not of programs! (asking if a program or a circuit is a member of \( \text{SIZE}_{n,m}(s) \) is a category error as in the sense of Fig. 4.11). As we discussed in Section 3.6.2 (and Section 2.5.1), the distinction between programs and functions is absolutely crucial. You should always remember that while a program computes a function, it is not equal to a function. In particular, as we’ve seen, there can be more than one program to compute the same function.

[There are \( 2^{2^n} \) functions mapping \( \{0, 1\}^n \) to \( \{0, 1\} \), and an infinite number of circuits with \( n \) bit inputs and a single bit of output. Every circuit computes one function, but every function can be computed by many circuits. We say that \( f \in \text{SIZE}_{n,1}(s) \) if the smallest circuit that computes \( f \) has \( s \) or fewer gates. For example \( \text{XOR}_n \in \text{SIZE}_{n,1}(4n) \). Theorem 4.12 shows that every function \( g \) is computable by some circuit of at most \( c \cdot 2^n/n \) gates, and hence \( \text{SIZE}_{n,1}(c \cdot 2^n/n) \) corresponds to the set of all functions from \( \{0, 1\}^n \) to \( \{0, 1\}^m \).]

While we defined \( \text{SIZE}(s) \) with respect to NAND gates, we would get essentially the same class if we defined it with respect to AND/OR/NOT gates:

**Lemma 4.19** Let \( \text{SIZE}^{\text{AO\&N}}_{n,m}(s) \) denote the set of all functions \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) that can be computed by an AND/OR/NOT Boolean circuit of at most \( s \) gates. Then,

\[
\text{SIZE}_{n,m}(s/2) \subseteq \text{SIZE}^{\text{AO\&N}}_{n,m}(s) \subseteq \text{SIZE}_{n,m}(3s) \quad (4.12)
\]

**Proof.** If \( f \) can be computed by a NAND circuit of at most \( s/2 \) gates, then by replacing each NAND with the two gates NOT and AND, we can obtain an AND/OR/NOT Boolean circuit of at most \( s \) gates that computes \( f \). On the other hand, if \( f \) can be computed by a Boolean
AND/OR/NOT circuit of at most \(s\) gates, then by Theorem 3.12 it can be computed by a NAND circuit of at most \(3s\) gates.

The results we have seen in this chapter can be phrased as showing that \(ADD_n \in \text{SIZE}_{2n,n+1}(100n)\) and \(MULT_n \in \text{SIZE}_{2n,2n}(10000n^{log_2 3})\). Theorem 4.12 shows that for some constant \(c\), \(\text{SIZE}_{n,m}(cm^{2^n})\) is equal to the set of all functions from \(\{0, 1\}^n\) to \(\{0, 1\}^m\).

**Remark 4.20 — Finite vs infinite functions.** A NAND-CIRC program \(P\) can only compute a function with a certain number \(n\) of inputs and a certain number \(m\) of outputs. Hence for example there is no single NAND-CIRC program that can compute the increment function \(\text{INC} : \{0, 1\}^n \to \{0, 1\}^n\) that maps a string \(x\) (which we identify with a number via the binary representation) to the string that represents \(x + 1\). Rather for every \(n > 0\), there is a NAND-CIRC program \(P_n\) that computes the restriction \(\text{INC}_n\) of the function \(\text{INC}\) to inputs of length \(n\). Since it can be shown that for every \(n > 0\) such a program \(P_n\) exists of length at most \(10n\), \(\text{INC}_n \in \text{SIZE}(10n)\) for every \(n > 0\).

If \(T : \mathbb{N} \to \mathbb{N}\) and \(F : \{0, 1\}^* \to \{0, 1\}^*\), we will write \(F \in \text{SIZE}(T(n))\) (or sometimes slightly abuse notation and write simply \(F \in \text{SIZE}(T(n))\)) to indicate that for every \(n\) the restriction \(F_n\) of \(F\) to inputs in \(\{0, 1\}^n\) is in \(\text{SIZE}(T(n))\). Hence we can write \(\text{INC} \in \text{SIZE}_{\log(10n)}\). We will come back to this issue of finite vs infinite functions later in this course.

**Solved Exercise 4.1 — \text{SIZE} closed under complement.** In this exercise we prove a certain “closure property” of the class \(\text{SIZE}_n(s)\). That is, we show that if \(f\) is in this class then (up to some small additive term) so is the complement of \(f\), which is the function \(g(x) = 1 - f(x)\).

Prove that there is a constant \(c\) such that for every \(f : \{0, 1\}^n \to \{0, 1\}\) and \(s \in \mathbb{N}\), if \(f \in \text{SIZE}_n(s)\) then \(1 - f \in \text{SIZE}_n(s + c)\).

**Solution:**

If \(f \in \text{SIZE}(s)\) then there is an \(s\)-line NAND-CIRC program \(P\) that computes \(f\). We can rename the variable \(Y[0]\) in \(P\) to a variable \(\text{temp}\) and add the line

\[Y[0] = \text{NAND}(\text{temp}, \text{temp})\]

at the very end to obtain a program \(P'\) that computes \(1 - f\).
Lecture Recap

- We can define the notion of computing a function via a simplified “programming language”, where computing a function \( F \) in \( T \) steps would correspond to having a \( T \)-line NAND-CIRC program that computes \( F \).
- While the NAND-CIRC programming only has one operation, other operations such as functions and conditional execution can be implemented using it.
- Every function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be computed by a circuit of at most \( O(m2^n) \) gates (and in fact at most \( O(m2^n/n) \) gates).
- Sometimes (or maybe always?) we can translate an efficient algorithm to compute \( f \) into a circuit that computes \( f \) with a number of gates comparable to the number of steps in this algorithm.

4.7 EXERCISES

Exercise 4.1 — Pairing. This exercise asks you to give a one-to-one map from \( \mathbb{N}^2 \) to \( \mathbb{N} \). This can be useful to implement two-dimensional arrays as “syntactic sugar” in programming languages that only have one-dimensional array.

1. Prove that the map \( F(x, y) = 2^x3^y \) is a one-to-one map from \( \mathbb{N}^2 \) to \( \mathbb{N} \).

2. Show that there is a one-to-one map \( F : \mathbb{N}^2 \rightarrow \mathbb{N} \) such that for every \( x, y \), \( F(x, y) \leq 100 \cdot \max\{x, y\}^2 + 100 \).

3. For every \( k \), show that there is a one-to-one map \( F : \mathbb{N}^k \rightarrow \mathbb{N} \) such that for every \( x_0, \ldots, x_{k-1} \in \mathbb{N} \), \( F(x_0, \ldots, x_{k-1}) \leq 100 \cdot (x_0 + x_1 + \ldots + x_{k-1} + 100k)^k \).

Exercise 4.2 — Computing MUX. Prove that the NAND-CIRC program below computes the function MUX (or \( LOOKUP_1 \)) where MUX\( (a, b, c) \) equals \( a \) if \( c = 0 \) and equals \( b \) if \( c = 1 \):

\[
\begin{align*}
t &= \text{NAND}(X[2], X[2]) \\
u &= \text{NAND}(X[0], t) \\
v &= \text{NAND}(X[1], X[2]) \\
Y[0] &= \text{NAND}(u, v)
\end{align*}
\]

Exercise 4.3 — At least two / Majority. Give a NAND-CIRC program of at most 6 lines to compute the function MAJ : \( \{0, 1\}^3 \rightarrow \{0, 1\} \) where \( \text{MAJ}(a, b, c) = 1 \) if \( a + b + c \geq 2 \).
Exercise 4.4 — Conditional statements. In this exercise we will explore Theorem 4.6: transforming NAND-CIRC-IF programs that use code such as if .. then .. else .. to standard NAND-CIRC programs.

1. Give a “proof by code” of Theorem 4.6: a program in a programming language of your choice that transforms a NAND-CIRC-IF program $P$ into a “sugar free” NAND-CIRC program $P'$ that computes the same function. See footnote for hint.6

2. Prove the following statement, which is the heart of Theorem 4.6: suppose that there exists an $s$-line NAND-CIRC program to compute $f : \{0, 1\}^n \rightarrow \{0, 1\}$ and an $s'$-line NAND-CIRC program to compute $g : \{0, 1\}^n \rightarrow \{0, 1\}$. Prove that there exist a NAND-CIRC program of at most $s + s' + 10$ lines to compute the function $h : \{0, 1\}^{n+1} \rightarrow \{0, 1\}$ where $h(x_0, \ldots, x_{n-1}, x_n)$ equals $f(x_0, \ldots, x_{n-1})$ if $x_n = 0$ and equals $g(x_0, \ldots, x_{n-1})$ otherwise. (All programs in this item are standard “sugar-free” NAND-CIRC programs.)

Exercise 4.5 — Half and full adders. 1. A half adder is the function $HA : \{0, 1\}^2 \rightarrow \{0, 1\}^2$ that corresponds to adding two binary bits. That is, for every $a, b \in \{0, 1\}$, $HA(a, b) = (e, f)$ where $2e + f = a + b$. Prove that there is a NAND circuit of at most five NAND gates that computes $HA$.

2. A full adder is the function $FA : \{0, 1\}^3 \rightarrow \{0, 1\}$ that takes in two bits and a “carry” bit and outputs their sum. That is, for every $a, b, c \in \{0, 1\}$, $FA(a, b, c) = (e, f)\$ such that $2e + f = a + b + c$. Prove that there is a NAND circuit of at most nine NAND gates that computes $FA$.

3. Prove that if there is a NAND circuit of $c$ gates that computes $FA$, then there is a circuit of $cn$ gates that computes $ADD_n$ where (as in Theorem 4.7) $ADD_n : \{0, 1\}^{2n} \rightarrow \{0, 1\}^n$ is the function that outputs the addition of two input $n$-bit numbers. See footnote for hint.7

4. Show that for every $n$ there is a NAND-CIRC program to compute $ADD_n$ with at most $9n$ lines.

Exercise 4.6 — Addition. Write a program using your favorite programming language that on input an integer $n$, outputs a NAND-CIRC program that computes $ADD_n$. Can you ensure that the program it outputs for $ADD_n$ has fewer than $10n$ lines?

---

6 You can start by transforming $P$ into a NAND-CIRC-PROC program that uses procedure statements, and then use the code of ?? to transform the latter into a “sugar free” NAND-CIRC program.

7 Use a “cascade” of adding the bits one after the other, starting with the least significant digit, just like in the elementary-school algorithm.
Exercise 4.7 — Multiplication. Write a program using your favorite programming language that on input an integer $n$, outputs a NAND-CIRC program that computes $MULT_n$. Can you ensure that the program it outputs for $MULT_n$ has fewer than $1000 \cdot n^2$ lines?

Exercise 4.8 — Efficient multiplication (challenge). Write a program using your favorite programming language that on input an integer $n$, outputs a NAND-CIRC program that computes $MULT_n$ and has at most $10000n^{1.9}$ lines. What is the smallest number of lines you can use to multiply two 2048 bit numbers? \footnote{Hint: Use Karatsuba’s algorithm.}

Exercise 4.9 — Multibit function. In the text Theorem 4.12 is only proven for the case $m = 1$. In this exercise you will extend the proof for every $m$.

Prove that

1. If there is an $s$-line NAND-CIRC program to compute $f : \{0, 1\}^n \rightarrow \{0, 1\}$ and an $s'$-line NAND-CIRC program to compute $f' : \{0, 1\}^n \rightarrow \{0, 1\}$ then there is an $s + s'$-line program to compute the function $g : \{0, 1\}^n \rightarrow \{0, 1\}^2$ such that $g(x) = (f(x), f'(x))$.

2. For every function $f : \{0, 1\}^n \rightarrow \{0, 1\}^m$, there is a NAND-CIRC program of at most $10m \cdot 2^n$ lines that computes $f$. (You can use the $m = 1$ case of Theorem 4.12, as well as Item 1.)

Exercise 4.10 — Simplifying using syntactic sugar. Let $P$ be the following NAND-CIRC program:

\[
\begin{align*}
\text{Temp}[0] &= \text{NAND}(X[0], X[0]) \\
\text{Temp}[1] &= \text{NAND}(X[1], X[1]) \\
\text{Temp}[2] &= \text{NAND}(\text{Temp}[0], \text{Temp}[1]) \\
\text{Temp}[3] &= \text{NAND}(X[2], X[2]) \\
\text{Temp}[4] &= \text{NAND}(X[3], X[3]) \\
\text{Temp}[5] &= \text{NAND}(\text{Temp}[3], \text{Temp}[4]) \\
\text{Temp}[6] &= \text{NAND}(\text{Temp}[2], \text{Temp}[2]) \\
\text{Temp}[7] &= \text{NAND}(\text{Temp}[5], \text{Temp}[5]) \\
Y[0] &= \text{NAND}(\text{Temp}[6], \text{Temp}[7])
\end{align*}
\]

1. Write a program $P'$ with at most three lines of code that uses both NAND as well as the syntactic sugar OR that computes the same function as $P$. 
2. Draw a circuit that computes the same function as $P$ and uses only
$\textit{AND}$ and $\textit{NOT}$ gates.

In the following exercises you are asked to compare the power of
pairs programming languages. By “comparing the power” of two
programming languages $X$ and $Y$ we mean determining the relation
between the set of functions that are computable using programs in $X$
and $Y$ respectively. That is, to answer such a question you need to do
both of:

1. Either prove that for every program $P$ in $X$ there is a program $P'$
in $Y$ that computes the same function as $P$, or give an example for
a function that is computable by an $X$-program but not computable
by a $Y$-program.

and

1. Either prove that for every program $P$ in $Y$ there is a program $P'$
in $X$ that computes the same function as $P$, or give an example for
a function that is computable by a $Y$-program but not computable by
an $X$-program.

When you give an example as above of a function that is com-
putable in one programming language but not the other, you need to
prove that the function you showed is (1) computable in the first pro-
gramming language, (2) not computable in the second programming
language.

Exercise 4.11 — Compare $\textit{IF}$ and $\textit{NAND}$. Let $\text{IF-CIRC}$ be the programming
language where we have the following operations $\texttt{foo} = 0$, $\texttt{foo} = 1$,
$\texttt{foo} = \texttt{IF}(\texttt{cond}, \texttt{yes}, \texttt{no})$ (that is, we can use the constants 0 and 1,
and the $\texttt{IF} : \{0,1\}^3 \rightarrow \{0,1\}$ function such that $\texttt{IF}(a, b, c)$ equals $b$ if
$a = 1$ and equals $c$ if $a = 0$). Compare the power of the $\textit{NAND}$-$\text{CIRC}$
programming language and the $\textit{IF}$-$\text{CIRC}$ programming language.

Exercise 4.12 — Compare $\textit{XOR}$ and $\textit{NAND}$. Let $\text{XOR-CIRC}$ be the pro-
gramming language where we have the following operations $\texttt{foo} = \texttt{XOR}(\texttt{bar}, \texttt{blah})$, $\texttt{foo} = 1$ and $\texttt{bar} = 0$ (that is, we can use the
constants 0, 1 and the $\texttt{XOR}$ function that maps $a, b \in \{0,1\}^2$ to $a + b$
mod 2). Compare the power of the $\textit{NAND}$-$\text{CIRC}$ programming
language and the $\textit{XOR}$-$\text{CIRC}$ programming language. See footnote for
hint.\(^9\)

Exercise 4.13 — Circuits for majority. Prove that there is some constant $c$
such that for every $n > 1$, $\text{MAJ}_n \in \text{Size}(cn)$ where $\text{MAJ}_n : \{0,1\}^n \rightarrow$

\(^9\) You can use the fact that $(a+b)+c \mod 2 = a+b+c \mod 2$. In particular it means that if you have the
lines $d = \text{XOR}(a, b)$ and $e = \text{XOR}(d, c)$ then $e$ gets the
sum modulo 2 of the variable $a, b$ and $c$.\)
\{0, 1\} is the majority function on \( n \) input bits. That is \( \text{MAJ}_n(x) = 1 \) iff \( \sum_{i=0}^{n-1} x_i > n/2 \). See footnote for hint.\(^\text{10}\)

**Exercise 4.14 — Circuits for threshold.** Prove that there is some constant \( c \) such that for every \( n > 1 \), and integers \( a_0, \ldots, a_{n-1}, b \in \{-2^n, -2^n + 1, \ldots, -1, 0, +1, \ldots, 2^n\} \), there is a NAND circuit with at most \( n^c \) gates that computes the threshold function \( f_{a_0, \ldots, a_{n-1}, b} : \{0, 1\}^n \to \{0, 1\} \) that on input \( x \in \{0, 1\}^n \) outputs 1 if and only if \( \sum_{i=0}^{n-1} a_i x_i > b \).

### 4.8 BIBLIOGRAPHICAL NOTES

See Jukna’s and Wegener’s books [Juk12; Weg87] for much more extensive discussion on circuits. Shannon showed that every Boolean function can be computed by a circuit of exponential size [Sha38]. The improved bound of \( c \cdot 2^n / n \) (with the optimal value of \( c \) for many bases) is due to Lupanov [Lup58]. An exposition of this for the case of NAND (where \( c = 1 \)) is given in Chapter 4 of his book [Lup84]. (Thanks to Sasha Golovnev for tracking down this reference!)

The concept of “syntactic sugar” is also known as “macros” or “meta-programming” and is sometimes implemented via a preprocessor or macro language in a programming language or a text editor. One modern example is the Babel JavaScript syntax transformer, that converts JavaScript programs written using the latest features into a format that older Browsers can accept. It even has a plug-in architecture, that allows users to add their own syntactic sugar to the language.

\(^\text{10}\) One approach to solve this is using recursion and the so-called Master Theorem.
5

**Code as data, data as code**

“*The term code script is, of course, too narrow. The chromosomal structures are at the same time instrumental in bringing about the development they foreshadow. They are law-code and executive power - or, to use another simile, they are architect’s plan and builder’s craft - in one.*”, Erwin Schrödinger, 1944.

“A mathematician would hardly call a correspondence between the set of 64 triples of four units and a set of twenty other units, “universal”, while such correspondence is, probably, the most fundamental general feature of life on Earth”, Misha Gromov, 2013

A program is simply a sequence of symbols, each of which can be encoded as a string of 0’s and 1’s using (for example) the ASCII standard. Therefore we can represent every NAND-CIRC program (and hence also every Boolean circuit) as a binary string. This statement seems obvious but it is actually quite profound. It means that we can treat circuits or NAND-CIRC programs both as instructions to carrying computation and also as data that could potentially be used as inputs to other computations.

☐ **Big Idea 6** A program is a piece of text, and so it can be fed as input to other programs.

This correspondence between code and data is one of the most fundamental aspects of computing. It underlies the notion of general purpose computers, that are not pre-wired to compute only one task, and also forms the basis of our hope for obtaining general artificial intelligence. This concept finds immense use in all areas of computing, from scripting languages to machine learning, but it is fair to say that we haven’t yet fully mastered it. Many security exploits involve cases such
as “buffer overflows” when attackers manage to inject code where the system expected only “passive” data (see Fig. 5.1). The relation between code and data reaches beyond the realm of electronic computers. For example, DNA can be thought of as both a program and data (in the words of Schrödinger, who wrote before DNA’s discovery a book that inspired Watson and Crick, DNA is both “architect’s plan and builder’s craft”).

In this chapter, we will begin to explore some of the applications of this connection. We start by using the representation of programs/circuits as strings to count the number of programs/circuits up to a certain size, and use that to obtain a counterpart to the result we proved in Chapter 4. There we proved that for every function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \), there exists a circuit of at most \( 100 \cdot 2^n/n \) gates to compute it. (The number 100 here is somewhat arbitrary and fixed for concreteness; Theorem 4.16 states a bound of \( c \cdot 2^n/n \) for some constant \( c \), but it can be verified that the proof yields \( c \leq 100 \).) In this chapter we will prove that there are some functions \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) for which we cannot do much better: they require a circuit of size at least \( 0.01 \cdot 2^n/n \) (see Theorem 5.3). We will also use the notion of representing programs/circuits as strings to show the existence of a bounded universal circuit \( U \) that gets as input the string representation of another circuit \( C \) and a string \( x \), and outputs \( C(x) \). (The qualifier “bounded” means that the circuit \( C \) has to be of at most a certain size; we see computational models that overcome this limitation in Chapter 6, which introduces the notion of programming languages with loops and the computational model of a Turing Machine.) Equivalently, taking the programming-language point of view, the bounded universal circuit corresponds to a “NAND-CIRC interpreter in NAND-CIRC”: a NAND-CIRC program that can evaluate other NAND-CIRC program. Such a program is known in Computer Science as a “meta-circular evaluator”, and is fundamental to both theory and practice of computing. See Fig. 5.2 for an overview of the results of this chapter.

### 5.1 REPRESENTING PROGRAMS AS STRINGS
We can represent programs or circuits as strings in a myriad of ways. For example, since Boolean circuits are labeled directed acyclic graphs, we can use the adjacency matrix or adjacency list representations for them. However, since the code of a program is ultimately just a sequence of letters and symbols, arguably the conceptually simplest representation of a program is as such a sequence. For example, the following NAND-CIRC program $P$

\[
\begin{align*}
temp_0 &= \text{NAND}(X[0],X[1]) \\
temp_1 &= \text{NAND}(X[0],\text{temp}_0) \\
temp_2 &= \text{NAND}(X[1],\text{temp}_0) \\
Y[0] &= \text{NAND}(\text{temp}_1,\text{temp}_2)
\end{align*}
\]

is simply a string of 107 symbols which include lower and upper case letters, digits, the underscore character \_ and equality sign =, punctuation marks such as ‘(‘, ‘)’, ‘,’ , ‘,’ spaces, and “new line” markers (often denoted as \textbackslash n or \textbackslash \textbackslash a). Each such symbol can be encoded as a string of 7 bits using the ASCII encoding, and hence the program $P$ can be encoded as a string of length $7 \cdot 107 = 749$ bits.

Nothing in the above discussion was specific to the program $P$, and hence we can use the same reasoning to prove that every NAND-CIRC program can be represented as a string in $\{0, 1\}^*$. In fact, we can do a bit better. Since the names of the working variables of a NAND-CIRC program do not effect its functionality, we can always transform a program to have the form of $P'$ where all variables apart from the inputs and outputs have the form $\text{temp}_0, \text{temp}_1, \text{temp}_2, \text{etc.}$. Moreover, if the program has $s$ lines, then we will never need to use an index larger than $3s$ (since each line involves at most three variables), and similarly the indices of the input and output variables will all be at most $3s$. Since a number between 0 and $3s$ can be expressed using at most $\lceil \log_{10}(3s + 1) \rceil = O(\log s)$ digits, each line in the program (which has the form $\text{foo} = \text{NAND}(\text{bar},\text{blah})$), can be represented

**Figure 5.2**: Overview of the results in this chapter. We use the representation of programs/circuits as strings to derive two main results. First we show the existence of a universal program/circuit, and in fact (with more work) the existence of such a program/circuit whose size is at most polynomial in the size of the program/circuit it evaluates. We then use the string representation to count the number of programs/circuits of a given size, and use that to establish that some functions require an exponential number of lines/gates to compute.

**Figure 5.3**: In the Harvard Mark I computer, a program was represented as a list of triples of numbers, which were then encoded by perforating holes in a control card.
using \(O(1) + O(\log s) = O(\log s)\) symbols, each of which can be represented by 7 bits. Hence an \(s\) line program can be represented as a string of \(O(s \log s)\) bits, resulting in the following theorem:

**Theorem 5.1 — Representing programs as strings.** There is a constant \(c\) such that for \(f \in \text{SIZE}(s)\), there exists a program \(P\) computing \(f\) that whose string representation has length at most \(cs \log s\).

We omit the formal proof of Theorem 5.1 but please make sure that you understand why it follows from the reasoning above.

### 5.2 COUNTING PROGRAMS, AND LOWER BOUNDS ON THE SIZE OF NAND-CIRC PROGRAMS

One consequence of the representation of programs as strings is that the number of programs of certain length is bounded by the number of strings that represent them. This has consequences for the sets \(\text{SIZE}(s)\) that we defined in Section 4.6.

**Theorem 5.2 — Counting programs.** For every \(s \in \mathbb{N}\),

\[
|\text{SIZE}(s)| \leq 2^{O(s \log s)}.
\]  

(5.1)

That is, there are at most \(2^{O(s \log s)}\) functions computed by NAND-CIRC programs of at most \(s\) lines. \(^1\)

**Proof.** We will show a one-to-one map \(E\) from \(\text{SIZE}(s)\) to the set of strings of length \(cs \log s\) for some constant \(c\). This will conclude the proof, since it implies that \(|\text{SIZE}(s)|\) is smaller than the size of the set of all strings of length at most \(\ell\), which equals \(1 + 2 + 4 + \cdots + 2^\ell = 2^{\ell+1} - 1\) by the formula for sums of geometric progressions.

The map \(E\) will simply map \(f\) to the representation of the program computing \(f\). Specifically, we let \(E(f)\) be the representation of the program \(P\) computing \(f\) given by Theorem 5.1. This representation has size at most \(cs \log s\), and moreover the map \(E\) is one to one, since if \(f \neq f'\) then every two programs computing \(f\) and \(f'\) respectively must have different representations.

\[ \blacksquare \]

A function mapping \(\{0, 1\}^2\) to \(\{0, 1\}\) can be identified with the table of its four values on the inputs 00, 01, 10, 11. A function mapping \(\{0, 1\}^3\) to \(\{0, 1\}\) can be identified with the table of its eight values on

\(^1\) The implicit constant in the \(O(\cdot)\) notation is smaller than 10. That is, for all sufficiently large \(s\), \(|\text{SIZE}(s)| < 2^{10s \log s}\), see Remark 5.4. As discussed in ??, we use the bound 10 simply because it is a round number.
the inputs 000, 001, 010, 011, 100, 101, 110, 111. More generally, every function \( F : \{0, 1\}^n \rightarrow \{0, 1\} \) can be identified with the table of its \( 2^n \) values on the inputs \( \{0, 1\}^n \). Hence the number of functions mapping \( \{0, 1\}^n \) to \( \{0, 1\} \) is equal to the number of such tables which (since we can choose either 0 or 1 for every row) is exactly \( 2^{2^n} \). Note that this is double exponential in \( n \), and hence even for small values of \( n \) (e.g., \( n = 10 \)) the number of functions from \( \{0, 1\}^n \) to \( \{0, 1\} \) is truly astronomical.\(^2\) This has the following important corollary:

**Theorem 5.3 — Counting argument lower bound.** There is a constant \( \delta > 0 \), such that for every sufficiently large \( n \), there is a function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) such that \( f \notin \text{SIZE} \left( \frac{\delta 2^n}{n} \right) \). That is, the shortest NAND-CIRC program to compute \( F \) requires at least \( \delta \cdot 2^n / n \) lines.

**Proof.** The proof is simple. If we let \( c \) be the constant such that \( |\text{SIZE}(s)| \leq 2^s \log s \) and \( \delta = 1/c \), then setting \( s = \delta 2^n / n \) we see that

\[
|\text{SIZE}(\frac{\delta 2^n}{n})| \leq 2^{\delta 2^n} \log s < 2^{c \delta 2^n} = 2^{2^n}
\]  

(5.2)

using the fact that since \( s < 2^n \), \( \log s < n \). But since \( |\text{SIZE}(s)| \) is smaller than the total number of functions mapping \( n \) bits to 1 bit, there must be at least one such function not in \( \text{SIZE}(s) \), which is what we needed to prove.

We have seen before that every function mapping \( \{0, 1\}^n \) to \( \{0, 1\} \) can be computed by an \( O(2^n / n) \) line program. Theorem 5.3 shows that this is tight in the sense that some functions do require such an astronomical number of lines to compute.

**Big Idea 7** Some functions \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) cannot be computed by a Boolean circuit using a fewer than exponential number of gates.

In fact, as we explore in the exercises, this is the case for most functions. Hence functions that can be computed in a small number of lines (such as addition, multiplication, finding short paths in graphs, or even the \textit{EVAL} function) are the exception, rather than the rule.
5.2.1 Size hierarchy theorem (optional)

By Theorem 4.15 the class \( \text{SIZE}_n(10 \cdot 2^n/n) \) contains all functions from \( \{0, 1\}^n \) to \( \{0, 1\} \), while by Theorem 5.3, there is some function \( f : \{0, 1\}^n \to \{0, 1\} \) that is not contained in \( \text{SIZE}_n(0.1 \cdot 2^n/n) \). In other words, for every sufficiently large \( n \),

\[
\text{SIZE}_n(0.1 \cdot 2^n/n) \subsetneq \text{SIZE}_n\left(10 \cdot 2^n/n\right).
\]  

It turns out that we can use Theorem 5.3 to show a more general result: whenever we increase our “budget” of gates we can compute new functions.

**Theorem 5.5 — Size Hierarchy Theorem.** For every sufficiently large \( n \) and \( 10n < s < 0.12^n/n \),

\[
\text{SIZE}_n(s) \subsetneq \text{SIZE}_n(s + 10n).
\]  

**Proof Idea:**

To prove the theorem we need to find a function \( f : \{0, 1\}^n \to \{0, 1\} \) such that \( f \) can be computed by a circuit of \( s + 10n \) gates but it cannot be computed by a circuit of \( s \) gates. We will do so by coming up with a sequence of functions \( f_0, f_1, f_2, \ldots, f_N \) with the following properties:

1. \( f_0 \) can be computed by a circuit of at most \( 10n \) gates,
2. \( f_N \) cannot be computed by a circuit of \( 0.1 \cdot 2^n/n \) gates, and
3. for every \( i \in \{0, \ldots, N\} \), if \( f_i \) can be computed by a circuit of size \( s \), then \( f_{i+1} \) can be computed by a circuit of size at most \( s + 10n \). Together these properties imply that if we let \( i \) be the smallest number such that \( f_i \notin \text{SIZE}_n(s) \), then since \( f_{i+1} \in \text{SIZE}(s) \) it must hold that \( f_i \in \text{SIZE}(s + 10n) \) which is what we need to prove. See Fig. 5.4 for an illustration.

*Proof of Theorem 5.5. Let \( f^* : \{0, 1\}^n \to \{0, 1\} \) be the function (whose existence we are guaranteed by Theorem 5.3) such that \( f^* \notin \text{SIZE}_n(0.1 \cdot 2^n/n) \). We define the functions \( f_0, f_1, \ldots, f_{2^n} \) mapping \( \{0, 1\}^n \) to \( \{0, 1\} \) as follows. For every \( x \in \{0, 1\}^n \), if \( \text{lex}(x) \in \{0, 1, \ldots, 2^n - 1\} \) is \( x \)'s order in the lexicographical order then

\[
f_i(x) = \begin{cases} f^*(x) & \text{\text{lex}}(x) < i \\ 0 & \text{otherwise} \end{cases}.
\]  

Figure 5.4: We prove Theorem 5.5 by coming up with a list \( f_0, \ldots, f_{2^n} \) of functions such that \( f_0 \) is the all zero function, \( f_{2^n} \) is a function (obtained from Theorem 5.3) outside of \( \text{SIZE}(0.1 \cdot 2^n/n) \) and such that \( f_{i-1} \) and \( f_i \) differ by one another on at most one input. We can show that for every \( i \), the number of gates to compute \( f_i \) is most \( 10n \) larger than the number of gates to compute \( f_{i-1} \) and so if let \( i \) be the smallest number such that \( f_i \notin \text{SIZE}(s) \), then \( f_i \in \text{SIZE}(s + 10n) \).
The function \( f_0 \) is simply the constant zero function, while the function \( f_{2^n} \) is equal to \( f^* \). Moreover, for every \( i \in [2^n] \), the function \( f_i \) and \( f_{i+1} \) differ on at most one input (i.e., the input \( x \in \{0, 1\}^n \) such that \( \text{lex}(x) = i \)). Let \( 10n < s < 0.1 \cdot 2^n / n \), and let \( i \) be the first index such that \( f_i \notin \text{SIZE}_n(s) \). Since \( f_{2^n} = f^* \notin \text{SIZE}(0.12^n / n) \) there must exist such an index \( i \), and moreover \( i > 0 \) since the constant zero function is a member of \( \text{SIZE}_n(10n) \).

By our choice of \( i \), \( f_{i-1} \) is a member of \( \text{SIZE}_n(s) \). To complete the proof, we need to show that \( f_i \in \text{SIZE}_n(s + 10n) \). Let \( x^* \) be the string such that \( \text{lex}(x^*) = i \). Let \( b \in \{0, 1\} \) be the value \( f^*(x^*) \). Then we can define \( f_i \) also as follows

\[
 f_i(x) = \begin{cases} 
 b & x = x^* \\
 f_{i-1}(x) & x \neq x^* 
\end{cases}
\]  

(5.6)

or in other words

\[
 f_i(x) = f_{i-1}(x) \land \text{EQUAL}(x^*, x) \lor b \land \neg \text{EQUAL}(x^*, x)
\]  

(5.7)

where \( \text{EQUAL} : \{0, 1\}^{2n} \rightarrow \{0, 1\} \) is the function that maps \( x, x' \in \{0, 1\}^n \) to 1 if they are equal and to 0 otherwise. Since (by our choice of \( i \)), \( f_{i-1} \) can be computed using at most \( s \) gates and (as can be easily verified) that \( \text{EQUAL} \in \text{SIZE}_n(9n) \), we can compute \( f_i \) using at most \( s + 9n + O(1) \leq s + 10n \) gates which is what we wanted to prove.

\[\Box\]

**Figure 5.5:** An illustration of some of what we know about the size complexity classes (not to scale!). This figure depicts classes of the form \( \text{SIZE}_{n,1}(s) \) but the state of affairs for other size complexity classes such as \( \text{SIZE}_{n,1}(s) \) is similar. We know by Theorem 4.12 (with the improvement of Section 4.4.2) that all functions mapping \( n \) bits to \( n \) bits can be computed by a circuit of size \( c \cdot 2^n \) for \( c \leq 10 \), while on the other hand the counting lower bound (Theorem 5.3, see also Exercise 5.4) shows that some such functions will require \( 0.1 \cdot 2^n \), and the size hierarchy theorem (Theorem 5.5) shows the existence of functions in \( \text{SIZE}(S) \setminus \text{SIZE}(s) \) whenever \( s = o(S) \), see also Exercise 5.5. We also consider some specific examples: addition of two \( n/2 \) bit numbers can be done in \( O(n) \) lines, while we don’t know of such a program for multiplying two \( n \) bit numbers, though we do know it can be done in \( O(n^2) \) and in fact even better size. In the above \( \text{FACTORN} \) corresponds to the inverse problem of multiplying- finding the prime factorization of a given number. At the moment we do not know of any circuit a polynomial (or even sub-exponential) number of lines that can compute \( \text{FACTORN} \).

**Remark 5.6 — Explicit functions.** While the size hierarchy theorem guarantees that there exists some function that can be computed using, for example, \( n^2 \) gates, but not using \( 100n \) gates, we do not know of any explicit
example of such a function. While we suspect that integer multiplication is such an example, we do not have any proof that this is the case.

5.3 THE TUPLES REPRESENTATION

ASCII is a fine presentation of programs, but for some applications it is useful to have a more concrete representation of NAND-CIRC programs. In this section we describe a particular choice, that will be convenient for us later on. A NAND-CIRC program is simply a sequence of lines of the form

blah = \text{NAND}(baz,boo)

There is of course nothing special about the particular names we use for variables. Although they would be harder to read, we could write all our programs using only working variables such as temp_0, temp_1 etc. Therefore, our representation for NAND-CIRC programs ignores the actual names of the variables, and just associate a number with each variable. We encode a line of the program as a triple of numbers. If the line has the form foo = \text{NAND}(bar,blah) then we encode it with the triple \((i,j,k)\) where \(i\) is the number corresponding to the variable foo and \(j\) and \(k\) are the numbers corresponding to bar and blah respectively.

More concretely, we use will associate every variable with a number in the set \([t] = \{0,1,\ldots,t-1\}\). The first \(n\) numbers \(\{0,\ldots,n-1\}\) correspond to the input variables, the last \(m\) numbers \(\{t-m,\ldots,t-1\}\) correspond to the output variables, and the intermediate numbers \(\{n,\ldots,t-m-1\}\) correspond to the remaining “workspace” variables. Formally, we define our representation as follows:

Definition 5.7 — List of tuples representation. Let \(P\) be a NAND-CIRC program of \(n\) inputs, \(m\) outputs, and \(s\) lines, and let \(t\) be the number of distinct variables used by \(P\). The list of tuples representation of \(P\) is the triple \((n,m,L)\) where \(L\) is a list of triples of the form \((i,j,k)\) for \(i,j,k \in [t]\).

We assign a number for variable of \(P\) as follows:

- For every \(i \in [n]\), the variable \(X[i]\) is assigned the number \(i\).
- For every \(j \in [m]\), the variable \(Y[j]\) is assigned the number \(t - m + j\).
- Every other variable is assigned a number in \(\{n,n+1,\ldots,t-m-1\}\) in the order in which the variable appears in the program \(P\).
The list of tuples representation is our default choice for representing NAND-CIRC programs. Since “list of tuples representation” is a bit of a mouthful, we will often call it simply “the representation” for a program $P$. Sometimes, when the number $n$ of inputs and number $m$ of outputs are known from the context, we will simply represent a program as the list $L$ instead of the triple $(n, m, L)$.

**Example 5.8 — Representing the XOR program.** Our favorite NAND-CIRC program, the program

\[
\begin{align*}
u &= \text{NAND}(X[0], X[1]) \\
v &= \text{NAND}(X[0], u) \\
w &= \text{NAND}(X[1], u) \\
Y[0] &= \text{NAND}(v, w)
\end{align*}
\]

computing the XOR function is represented as the tuple $(2, 1, L)$ where $L = ((2, 0, 1), (3, 0, 2), (4, 1, 2), (5, 3, 4))$. That is, the variables $X[0]$ and $X[1]$ are given the indices 0 and 1 respectively, the variables $u, v, w$ are given the indices 2, 3, 4 respectively, and the variable $Y[0]$ is given the index 5.

Transforming a NAND-CIRC program from its representation as code to the representation as a list of tuples is a fairly straightforward programming exercise, and in particular can be done in a few lines of Python.\(^4\) The list-of-tuples representation loses information such as the particular names we used for the variables, but this is OK since these names do not make a difference to the functionality of the program.

### 5.3.1 From tuples to strings

If $P$ is a program of size $s$, then the number $t$ of variables is at most $3s$ (as every line touches at most three variables). Hence we can encode every variable index in $[t]$ as a string of length $\ell = \lceil \log(3s) \rceil$, by adding leading zeroes as needed. Since this is a fixed-length encoding, it is prefix free, and so we can encode the list $L$ of $s$ triples (corresponding to the encoding of the $s$ lines of the program) as simply the string of length $3\ell s$ obtained by concatenating all of these encodings.

We define $S(s)$ to be the length of the string representing the list $L$ corresponding to a size $s$ program. By the above we see that

\[ S(s) = 3s\lceil \log(3s) \rceil. \tag{5.8} \]

We can represent $P = (n, m, L)$ as a string by prepending a prefix free representation of $n$ and $m$ to the list $L$. Since $n, m \leq 3s$ (a program must touch at least once all its input and output variables), those prefix free representations can be encoded using strings of length

\(^4\) If you’re curious what these few lines are, see our GitHub repository.
In particular, every program $P$ of at most $s$ lines can be represented by a string of length $O(s \log s)$. Similarly, every circuit $C$ of at most $s$ gates, can be represented by a string of length $O(s \log s)$ (for example by translating $C$ to the equivalent program $P$).

### 5.4 A NAND-CIRC INTERPRETER IN NAND-CIRC

Since we can represent programs as strings, we can also think of a program as an input to a function. In particular, for every natural numbers $s, n, m > 0$ we define the function $EVAL_{s,n,m} : \{0, 1\}^{S(s)+n} \rightarrow \{0, 1\}^m$ as follows:

$$EVAL_{s,n,m}(px) = \begin{cases} P(x) & p \in \{0, 1\}^{S(s)} \text{ represents a size-}s \text{ program } P \text{ with } n \text{ inputs and } m \text{ outputs} \\ 0^m & \text{otherwise} \end{cases}$$

(5.9)

where $S(s)$ is defined as in (5.8) and we use the concrete representation scheme described in Section 5.1.

That is, $EVAL_{s,n,m}$ takes as input the concatenation of two strings: a string $p \in \{0, 1\}^{S(s)}$ and a string $x \in \{0, 1\}^n$. If $p$ is a string that represents a list of triples $L$ such that $(n, m, L)$ is a list-of-tuples representation of a size-$s$ NAND-CIRC program $P$, then $EVAL_{s,n,m}(px)$ is equal to the evaluation $P(x)$ of the program $P$ on the input $x$. Otherwise, $EVAL_{s,n,m}(px)$ equals $0^m$ (this case is not very important: you can simply think of $0^m$ as some “junk value” that indicates an error).

**Take-away points.** The fine details of $EVAL_{s,n,m}$’s definition are not very crucial. Rather, what you need to remember about $EVAL_{s,n,m}$ is that:

- $EVAL_{s,n,m}$ is a finite function taking a string of fixed length as input and outputting a string of fixed length as output.

- $EVAL_{s,n,m}$ is a single function, such that computing $EVAL_{s,n,m}$ allows to evaluate arbitrary NAND-CIRC programs of a certain length on arbitrary inputs of the appropriate length.

- $EVAL_{s,n,m}$ is a function, not a program (recall the discussion in Section 3.6.2). That is, $EVAL_{s,n,m}$ is a specification of what output is associated with what input. The existence of a program that computes $EVAL_{s,n,m}$ (i.e., an implementation for $EVAL_{s,n,m}$) is a separate fact, which needs to be established (and which we will do in Theorem 5.9, with a more efficient program shown in in eff-bounded-univ).

One of the first examples of self circularity we will see in this book is the following theorem, which we can think of as showing a “NAND-CIRC interpreter in NAND-CIRC”:
That is, the NAND-CIRC program $U_{s,n,m}$ takes the description of any other NAND-CIRC program $P$ (of the right length and inputs/outputs) and any input $x$, and computes the result of evaluating the program $P$ on the input $x$. Given the equivalence between NAND-CIRC programs and Boolean circuits, we can also think of $U_{s,n,m}$ as a circuit that takes as input the description of other circuits and their inputs, and returns their evaluation, see Fig. 5.6. We call this NAND-CIRC program $U_{s,n,m}$ that computes $EVAL_{s,n,m}$ a bounded universal program (or a universal circuit, see Fig. 5.6). “Universal” stands for the fact that this is a single program that can evaluate arbitrary code, where “bounded” stands for the fact that $U_{s,n,m}$ only evaluates programs of bounded size. Of course this limitation is inherent for the NAND-CIRC programming language, since a program of $s$ lines (or, equivalently, a circuit of $s$ gates) can take at most $2s$ inputs. Later, in Chapter 6, we will introduce the concept of loops (and the model of Turing Machines), that allow to escape this limitation.

**Proof.** Theorem 5.9 is an important result, but it is actually not hard to prove. Specifically, since $EVAL_{s,n,m}$ is a finite function Theorem 5.9 is an immediate corollary of Theorem 4.12, which states that every finite function can be computed by some NAND-CIRC program.

---

**Theorem 5.9 — Bounded Universality of NAND-CIRC programs.** For every $s, n, m \in \mathbb{N}$ with $s \geq m$ there is a NAND-CIRC program $U_{s,n,m}$ that computes the function $EVAL_{s,n,m}$.  

---

**5.4.1 Efficient universal programs**

Theorem 5.9 establishes the existence of a NAND-CIRC program for computing $EVAL_{s,n,m}$, but it provides no explicit bound on the size of this program. Theorem 4.12, which we used to prove Theorem 5.9, guarantees the existence of a NAND-CIRC program whose size can be as large as exponential in the length of its input. This would mean that even for moderately small values of $s, n, m$ (for example $n = 100, s = 300, m = 1$), computing $EVAL_{s,n,m}$ might require a NAND program with more lines than there are atoms in the observable universe! Fortunately, we can do much better than that. In fact, for every $s, n, m$ there exists a NAND-CIRC program for comput-

---

**Figure 5.6**: A universal circuit $U$ is a circuit that gets as input the description of an arbitrary (smaller) circuit $P$ as a binary string, and an input $x$, and outputs the string $P(x)$ which is the evaluation of $P$ on $x$. We can also think of $U$ as a straight-line program that gets as input the code of a straight-line program $P$ and an input $x$, and outputs $P(x)$.  

---

Theorem 5.9 is simple but important. Make sure you understand what this theorem means, and why it is a corollary of Theorem 4.12.
ing $\text{EVAL}_{s,n,m}$ with size that is polynomial in its input length. This is shown in the following theorem.

**Theorem 5.10 — Efficient bounded universality of NAND-CIRC programs.**

For every $s, n, m \in \mathbb{N}$ there is a NAND-CIRC program of at most $O(s^2 \log s)$ lines that computes the function $\text{EVAL}_{s,n,m} : \{0, 1\}^{S+n} \to \{0, 1\}^m$ defined above (where $S$ is the number of bits needed to represent programs of $s$ lines).

If you haven’t done so already, now might be a good time to review $O$ notation in Section 1.4.8. In particular, an equivalent way to state Theorem 5.10 is that it says that there exists some number $c > 0$ such that for every $s, n, m \in \mathbb{N}$, there exists a NAND-CIRC program $P$ of at most $cs^2 \log s$ lines that computes the function $\text{EVAL}_{s,n,m}$.

Unlike Theorem 5.9, Theorem 5.10 is not a trivial corollary of the fact that every finite function can be computed by some circuit. Proving Theorem 5.9 requires us to present a concrete NAND-CIRC program for computing the function $\text{EVAL}_{s,n,m}$. We will do so in several stages.

1. First, we will describe the algorithm to evaluate $\text{EVAL}_{s,n,m}$ in “pseudo code”.
2. Then, we will show how we can write a program to compute $\text{EVAL}_{s,n,m}$ in Python. We will not use much about Python, and a reader that has familiarity with programming in any language should be able to follow along.
3. Finally, we will show how we can transform this Python program into a NAND-CIRC program.

This approach yields much more than just proving Theorem 5.10: we will see that it is in fact always possible to transform (loop free) code in high level languages such as Python to NAND-CIRC programs (and hence to Boolean circuits as well).

**5.4.2 A NAND-CIRC interpreter in “pseudocode”**

To prove Theorem 5.10 it suffices to give a NAND-CIRC program of $O(s^2 \log s)$ lines that can evaluate NAND-CIRC programs of $s$ lines. Let us start by thinking how we would evaluate such programs if we weren’t restricted to only performing NAND operations. That is, let us describe informally an *algorithm* that on input $n, m, s$, a list of triples...
$L$, and a string $x \in \{0, 1\}^n$, evaluates the program represented by $(n, m, L)$ on the string $x$.

It would be highly worthwhile for you to stop here and try to solve this problem yourself. For example, you can try thinking how you would write a program $\text{NADEVAL}(n, m, s, L, x)$ that computes this function in the programming language of your choice.

We will now describe such an algorithm. We assume that we have access to a bit array data structure that can store for every $i \in [t]$ a bit $T_i \in \{0, 1\}$. Specifically, if $\text{Table}$ is a variable holding this data structure, then we assume we can perform the operations:

- $\text{GET(\text{Table}, i)}$ which retrieves the bit corresponding to $i$ in $\text{Table}$. The value of $i$ is assumed to be an integer in $[t]$.

- $\text{Table} = \text{UPDATE(\text{Table}, i, b)}$ which updates $\text{Table}$ so the bit corresponding to $i$ is now set to $b$. The value of $i$ is assumed to be an integer in $[t]$ and $b$ is a bit in $\{0, 1\}$.

Algorithm 5.11 — Eval NAND-CIRC programs.

Input: Numbers $n, m, s$ and $t \leq 3s$, as well as a list $L$ of $s$ triples of numbers in $[t]$, and a string $x \in \{0, 1\}^n$.

Output: Evaluation of the program represented by $(n, m, L)$ on the Input: $x \in \{0, 1\}^n$.

1: Let $\text{Vartable}$ be table of size $t$
2: for $i$ in $[n]$ do
3: \hspace{1em} $\text{Vartable} = \text{UPDATE(\text{Vartable}, i, x_i)}$
4: end for
5: for $(i, j, k)$ in $L$ do
6: \hspace{1em} $a \leftarrow \text{GET(\text{Vartable}, j)}$
7: \hspace{1em} $b \leftarrow \text{GET(\text{Vartable}, k)}$
8: \hspace{1em} $\text{Vartable} = \text{UPDATE(\text{Vartable}, i, NAND(a, b))}$
9: end for
10: for $j$ in $[m]$ do
11: \hspace{1em} $y_j \leftarrow \text{GET(\text{Vartable}, t - m + j)}$
12: end for
13: return $y_0, \ldots, y_{m-1}$

Algorithm 5.11 evaluates the program given to it as input one line at a time, updating the $\text{Vartable}$ table to contain the value of each
variable. At the end of the execution it outputs the variables at positions \( t - m, t - m + 1, \ldots, t - 1 \) which correspond to the input variables.

5.4.3 A NAND interpreter in Python

To make things more concrete, let us see how we implement Algorithm 5.11 in the Python programming language. (There is nothing special about Python. We could have easily presented a corresponding function in JavaScript, C, OCaml, or any other programming language.) We will construct a function \( \text{NANDEVAL} \) that on input \( n, m, L, x \) will output the result of evaluating the program represented by \((n, m, L)\) on \( x \). To keep things simple, we will not worry about the case that \( L \) does not represent a valid program of \( n \) inputs and \( m \) outputs. The code is presented in Fig. 5.7.

Accessing an element of the array \( \text{Vartable} \) at a given index takes a constant number of basic operations. Hence (since \( n, m \leq s \) and \( t \leq 3s \)), the program above will use \( O(s) \) basic operations.\(^5\)

5.4.4 Constructing the NAND-CIRC interpreter in NAND-CIRC

We now turn to describing the proof of Theorem 5.10. To prove the theorem it is not enough to give a Python program. Rather, we need to show how we compute the function \( \text{EVAL}_{s,n,m} \) using a NAND-CIRC program. In other words, our job is to transform, for every \( s, n, m \), the Python code of Section 5.4.3 to a NAND-CIRC program \( U_{s,n,m} \) that computes the function \( \text{EVAL}_{s,n,m} \).

Before reading further, try to think how you could give a “constructive proof” of Theorem 5.10. That is, think of how you would write, in the programming language of your choice, a function \( \text{universal}(s,n,m) \) that on input \( s, n, m \) outputs the code for the NAND-CIRC program \( U_{s,n,m} \) such that \( U_{s,n,m} \) computes \( \text{EVAL}_{s,n,m} \). There is a subtle but crucial difference between this function and the Python \( \text{NANOEVAL} \) program described above. Rather than actually evaluating a given program \( P \) on some input \( w \), the function \( \text{universal} \) should output the code of a NAND-CIRC program that computes the map \( (P, x) \mapsto P(x) \).

Our construction will follow very closely the Python implementation of \( \text{EVAL} \) above. We will use variables \( \text{Vartable}[0], \ldots, \text{Vartable}[2^\ell - 1] \), where \( \ell = \lceil \log 3s \rceil \) to store our variables. However, NAND doesn’t have integer-valued variables, so we cannot write code such as \( \text{Vartable}[i] \) for some variable \( i \). However, we can implement the function \( \text{GET(Vartable, i)} \) that outputs the \( i \)-th bit of the array \( \text{Vartable} \). Indeed, this is nothing but the function \( \text{LOOKUP}_\ell \) that we have seen in Theorem 4.10!
def NAND_EVAL(n, m, L, X):
    # Evaluate a NAND-CIRC program from list of tuple representation.
    s = len(L)  # num of lines
    t = max(max(a, b, c) for (a, b, c) in L) + 1  # max index in L + 1
    Vartable = [0] * t  # initialize array

    # helper functions
    def GET(V, i): return V[i]
    def UPDATE(V, i, b):
        V[i] = b
        return V

    # load input values to Vartable:
    for i in range(n):
        Vartable = UPDATE(Vartable, i, X[i])

    # Run the program
    for (i, j, k) in L:
        a = GET(Vartable, j)
        b = GET(Vartable, k)
        c = NAND(a, b)
        Vartable = UPDATE(Vartable, i, c)

    # Return outputs Vartable[t-m], Vartable[t-m+1],...,Vartable[t-1]
    return [GET(Vartable, t-m+j) for j in range(m)]

# Test on XOR (2 inputs, 1 output)
L = ((2, 0, 1), (3, 0, 2), (4, 1, 2), (5, 3, 4))
print(NAND_EVAL(2, 1, L, (0, 1)))  # XOR(0,1)
# [1]
print(NAND_EVAL(2, 1, L, (1, 1)))  # XOR(1,1)
# [0]
Please make sure that you understand why GET and LOOKUP are the same function.

We saw that we can compute \( \text{LOOKUP}_\ell \) in time \( O(2^\ell) = O(s) \) for our choice of \( \ell \).

For every \( \ell \), let \( \text{UPDATE}_\ell : \{0,1\}^{2^\ell+\ell+1} \rightarrow \{0,1\}^{2^\ell} \) correspond to the UPDATE function for arrays of length \( 2^\ell \). That is, on input \( V \in \{0,1\}^{2^\ell} \), \( i \in \{0,1\}^\ell \), \( b \in \{0,1\} \), \( \text{UPDATE}_\ell(V, b, i) \) is equal to \( V' \in \{0,1\}^{2^\ell} \) such that

\[
V'_j = \begin{cases} 
V_j & j \neq i \\
b & j = 1 
\end{cases}
\]

(5.10)

where we identify the string \( i \in \{0,1\}^\ell \) with a number in \( \{0, \ldots, 2^\ell - 1\} \) using the binary representation. We can compute \( \text{UPDATE}_\ell \) using an \( O(2^{\ell+\ell}) = O(s \log s) \) line NAND-CIRC program as follows:

1. For every \( j \in [2^\ell] \), there is an \( O(\ell) \) line NAND-CIRC program to compute the function \( \text{EQUALS}_j : \{0,1\}^\ell \rightarrow \{0,1\} \) that on input \( i \) outputs 1 if and only if \( i \) is equal to (the binary representation of) \( j \). (We leave verifying this as Exercise 5.2 and Exercise 5.3.)

2. We have seen that we can compute the function \( \text{IF} : \{0,1\}^3 \rightarrow \{0,1\} \) such that \( \text{IF}(a, b, c) \) equals \( b \) if \( a = 1 \) and \( c \) if \( a = 0 \).

Together, this means that we can compute \( \text{UPDATE} \) (using some “syntactic sugar” for bounded length loops) as follows:

```python
def UPDATE_ell(V, i, b):
    # Get V[0]...V[2^ell-1], i in {0,1}^ell, b in {0,1}
    # Return NewV[0],...,NewV[2^ell-1]
    # updated array with NewV[i]=b and all
    # else same as V
    for j in range(2**ell):  # j = 0,1,2,...,2^ell -1
        a = EQUALS_j(i)
        NewV[j] = IF(a,b,V[j])
    return NewV
```

Since the loop over \( j \) in \( \text{UPDATE} \) is run \( 2^\ell \) times, and computing \( \text{EQUALS}_j \) takes \( O(\ell) \) lines, the total number of lines to compute \( \text{UPDATE} \) is \( O(2^\ell \cdot \ell) = O(s \log s) \). Once we can compute \( \text{GET} \) and \( \text{UPDATE} \), the rest of the implementation amounts to “book keeping” that needs to be done carefully, but is not too insightful, and hence we omit the full details. Since we run \( \text{GET} \) and \( \text{UPDATE} \) \( s \) times, the total number of lines for computing \( \text{EVAL}_{s,n,m} \) is \( O(s^2) + O(s^2 \log s) = O(s^2 \log s) \). This completes (up to the omitted details) the proof of Theorem 5.10.
5.5 A PYTHON INTERPRETER IN NAND-CIRC (DISCUSSION)

To prove Theorem 5.10 we essentially translated every line of the Python program for \texttt{EVAL} into an equivalent NAND-CIRC snip- pet. However none of our reasoning was specific to the particular function \texttt{EVAL}. It is possible to translate every Python program into an equivalent NAND-CIRC program of comparable efficiency. (More concretely, if the Python program takes \(T(n)\) operations on inputs of length at most \(n\) then there exists NAND-CIRC program of \(O(T(n) \log T(n))\) lines that agrees with the Python program on inputs of length \(n\).) Actually doing so requires taking care of many details and is beyond the scope of this book, but let me try to convince you why you should believe it is possible in principle.

For starters, one can can use \texttt{CPython} (the reference implementation for Python), to evaluate every Python program using a \texttt{C} program. We can combine this with a \texttt{C} compiler to transform a Python program to various flavors of “machine language”. So, to transform a Python program into an equivalent NAND-CIRC program, it is enough to show how to transform a \texttt{machine language} program into an equivalent NAND-CIRC program. One minimalistic (and hence
convenient) family of machine languages is known as the ARM architecture which powers many mobile devices including essentially all Android devices. There are even simpler machine languages, such as the LEG architecture for which a backend for the LLVM compiler was implemented (and hence can be the target of compiling any of large and growing list of languages that this compiler supports). Other examples include the TinyRAM architecture (motivated by interactive proof systems that we will discuss in Chapter 21) and the teaching-oriented Ridiculously Simple Computer architecture. Going one by one over the instruction sets of such computers and translating them to NAND snippets is no fun, but it is a feasible thing to do. In fact, ultimately this is very similar to the transformation that takes place in converting our high level code to actual silicon gates that are not so different from the operations of a NAND-CIRC program. Indeed, tools such as MyHDL that transform “Python to Silicon” can be used to convert a Python program to a NAND-CIRC program.

The NAND-CIRC programming language is just a teaching tool, and by no means do I suggest that writing NAND-CIRC programs, or compilers to NAND-CIRC, is a practical, useful, or enjoyable activity. What I do want is to make sure you understand why it can be done, and to have the confidence that if your life (or at least your grade) depended on it, then you would be able to do this. Understanding how programs in high level languages such as Python are eventually transformed into concrete low-level representation such as NAND is fundamental to computer science.

The astute reader might notice that the above paragraphs only outlined why it should be possible to find for every particular Python-computable function \( f \), a particular comparably efficient NAND-CIRC program \( P \) that computes \( f \). But this still seems to fall short of our goal of writing a “Python interpreter in NAND” which would mean that for every parameter \( n \), we come up with a single NAND-CIRC program \( UNIV_s \) such that given a description of a Python program \( P \), a particular input \( x \), and a bound \( T \) on the number of operations (where the lengths of \( P \) and \( x \) and the value of \( T \) are all at most \( s \)) returns the result of executing \( P \) on \( x \) for at most \( T \) steps. After all, the transformation above takes every Python program into a different NAND-CIRC program, and so does not yield “one NAND-CIRC program to rule them all” that can evaluate every Python program up to some given complexity. However, we can in fact obtain one NAND-CIRC program to evaluate arbitrary Python programs. The reason is that there exists a Python interpreter in Python: a Python program \( U \) that takes a bit string, interprets it as Python code, and then runs that code. Hence, we only need to show a NAND-CIRC program \( U^* \) that

\(^6\) ARM stands for “Advanced RISC Machine” where RISC in turn stands for “Reduced instruction set computer.”
computes the same function as the particular Python program \( U \), and this will give us a way to evaluate all Python programs.

What we are seeing time and again is the notion of universality or self reference of computation, which is the sense that all reasonably rich models of computation are expressive enough that they can “simulate themselves”. The importance of this phenomena to both the theory and practice of computing, as well as far beyond it, including the foundations of mathematics and basic questions in science, cannot be overstated.

5.6 THE PHYSICAL EXTENDED CHURCH-TURING THESIS (DISCUSSION)

We’ve seen that NAND gates (and other Boolean operations) can be implemented using very different systems in the physical world. What about the reverse direction? Can NAND-CIRC programs simulate any physical computer?

We can take a leap of faith and stipulate that Boolean circuits (or equivalently NAND-CIRC programs) do actually encapsulate every computation that we can think of. Such a statement (in the realm of infinite functions, which we’ll encounter in Chapter 6) is typically attributed to Alonzo Church and Alan Turing, and in that context is known as the Church Turing Thesis. As we will discuss in future lectures, the Church-Turing Thesis is not a mathematical theorem or conjecture. Rather, like theories in physics, the Church-Turing Thesis is about mathematically modeling the real world. In the context of finite functions, we can make the following informal hypothesis or prediction:

“Physical Extended Church-Turing Thesis”

(PECTT): If a function \( F : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be computed in the physical world using \( s \) amount of “physical resources” then it can be computed by a Boolean circuit program of roughly \( s \) gates.

A priori it might seem rather extreme to hypothesize that our meager model of NAND-CIRC programs or Boolean circuits captures all possible physical computation. But yet, in more than a century of computing technologies, no one has yet built any scalable computing device that challenges this hypothesis.

We now discuss the “fine print” of the PECTT in more detail, as well as the (so far unsuccessful) challenges that have been raised against it. There is no single universally-agreed-upon formalization of “roughly \( s \) physical resources”, but we can approximate this notion by considering the size of any physical computing device and the time it takes to compute the output, and ask that any such device can
be simulated by a Boolean circuit with a number of gates that is a polynomial (with not too large exponent) in the size of the system and the time it takes it to operate.

In other words, we can phrase the PECTT as stipulating that any function that can be computed by a device of that takes as a certain volume $V$ of space and requires $t$ time to complete the computation, must be computable by a Boolean circuit with a number of gates $p(V, t)$ that is polynomial in $V$ and $t$.

The exact form of the function $p(V, t)$ is not universally agreed upon but it is generally accepted that if $f : \{0, 1\}^n \to \{0, 1\}$ is an exponentially hard function, in the sense that it has no NAND-CIRC program of fewer than, say, $2^{n/2}$ lines, then a demonstration of a physical device that can compute in the real world $f$ for moderate input lengths (e.g., $n = 500$) would be a violation of the PECTT.

Remark 5.13 — Advanced note: making PECTT concrete (advanced, optional). We can attempt at a more exact phrasing of the PECTT as follows. Suppose that $Z$ is a physical system that accepts $n$ binary stimuli and has a binary output, and can be enclosed in a sphere of volume $V$. We say that the system $Z$ computes a function $f : \{0, 1\}^n \to \{0, 1\}$ within $t$ seconds if whenever we set the stimuli to some value $x \in \{0, 1\}^n$, if we measure the output after $t$ seconds then we obtain $f(x)$.

One can then phrase the PECTT as stipulating that if there exists such a system $Z$ that computes $F$ within $t$ seconds, then there exists a NAND-CIRC program that computes $F$ and has at most $\alpha(Vt)^2$ lines, where $\alpha$ is some normalization constant. (We can also consider variants where we use surface area instead of volume, or take $(Vt)$ to a different power than 2. However, none of these choices makes a qualitative difference to the discussion below.) In particular, suppose that $f : \{0, 1\}^n \to \{0, 1\}$ is a function that requires $2^n/(100n) > 2^{0.8n}$ lines for any NAND-CIRC program (such a function exists by Theorem 5.3). Then the PECTT would imply that either the volume or the time of a system that computes $F$ will have to be at least $2^{0.2n}/\sqrt{\alpha}$. Since this quantity grows exponentially in $n$, it is not hard to set parameters so that even for moderately large values of $n$, such a system could not fit in our universe.

To fully make the PECTT concrete, we need to decide on the units for measuring time and volume, and the normalization constant $\alpha$. One conservative choice is to assume that we could squeeze computation to the absolute physical limits (which are many orders of magnitude beyond current technology). This corresponds to setting $\alpha = 1$ and using the Planck units.
for volume and time. The Planck length $\ell_P$ (which is, roughly speaking, the shortest distance that can theoretically be measured) is roughly $2^{-120}$ meters. The Planck time $t_P$ (which is the time it takes for light to travel one Planck length) is about $2^{-150}$ seconds. In the above setting, if a function $F$ takes, say, 1KB of input (e.g., roughly $10^4$ bits, which can encode a 100 by 100 bitmap image), and requires at least $2^{0.8n} = 2^{0.8 \cdot 10^4}$ NAND lines to compute, then any physical system that computes it would require either volume of $2^{0.2 \cdot 10^4}$ Planck length cubed, which is more than $2^{1500}$ meters cubed or take at least $2^{0.2 \cdot 10^4}$ Planck Time units, which is larger than $2^{1500}$ seconds. To get a sense of how big that number is, note that the universe is only about $2^{100}$ seconds old, and its observable radius is only roughly $2^{200}$ meters. The above discussion suggests that it is possible to empirically falsify the PECTT by presenting a smaller-than-universe-size system that computes such a function.

There are of course several hurdles to refuting the PECTT in this way, one of which is that we can’t actually test the system on all possible inputs. However, it turns out that we can get around this issue using notions such as interactive proofs and program checking that we might encounter later in this book. Another, perhaps more salient problem, is that while we know many hard functions exist, at the moment there is no single explicit function $F : \{0, 1\}^n \to \{0, 1\}$ for which we can prove an $\omega(n)$ (let alone $\Omega(2^n/n)$) lower bound on the number of lines that a NAND-CIRC program needs to compute it.

### 5.6.1 Attempts at refuting the PECTT

One of the admirable traits of mankind is the refusal to accept limitations. In the best case this is manifested by people achieving long-standing “impossible” challenges such as heavier-than-air flight, putting a person on the moon, circumnavigating the globe, or even resolving Fermat’s Last Theorem. In the worst case it is manifested by people continually following the footsteps of previous failures to try to do proven-impossible tasks such as build a perpetual motion machine, trisect an angle with a compass and straightedge, or refute Bell’s inequality. The Physical Extended Church Turing thesis (in its various forms) has attracted both types of people. Here are some physical devices that have been speculated to achieve computational tasks that cannot be done by not-too-large NAND-CIRC programs:

- **Spaghetti sort:** One of the first lower bounds that Computer Science students encounter is that sorting $n$ numbers requires making $\Omega(n \log n)$ comparisons. The “spaghetti sort” is a description of a
proposed “mechanical computer” that would do this faster. The idea is that to sort \( n \) numbers \( x_1, \ldots, x_n \), we could cut \( n \) spaghetti noodles into lengths \( x_1, \ldots, x_n \), and then if we simply hold them together in our hand and bring them down to a flat surface, they will emerge in sorted order. There are a great many reasons why this is not truly a challenge to the PECTT hypothesis, and I will not ruin the reader’s fun in finding them out by her or himself.

- **Soap bubbles**: One function \( F : \{0, 1\}^n \to \{0, 1\} \) that is conjectured to require a large number of NAND lines to solve is the *Euclidean Steiner Tree* problem. This is the problem where one is given \( m \) points in the plane \((x_1, y_1), \ldots, (x_m, y_m)\) (say with integer coordinates ranging from 1 till \( m \), and hence the list can be represented as a string of \( n = O(m \log m) \) size) and some number \( K \). The goal is to figure out whether it is possible to connect all the points by line segments of total length at most \( K \). This function is conjectured to be hard because it is NP complete - a concept that we’ll encounter later in this course - and it is in fact reasonable to conjecture that as \( m \) grows, the number of NAND lines required to compute this function grows exponentially in \( m \), meaning that the PECTT would predict that if \( m \) is sufficiently large (such as few hundreds or so) then no physical device could compute \( F \). Yet, some people claimed that there is in fact a very simple physical device that could solve this problem, that can be constructed using some wooden pegs and soap. The idea is that if we take two glass plates, and put \( m \) wooden pegs between them in the locations \((x_1, y_1), \ldots, (x_m, y_m)\) then bubbles will form whose edges touch those pegs in the way that will minimize the total energy which turns out to be a function of the total length of the line segments. The problem with this device of course is that nature, just like people, often gets stuck in “local optima”. That is, the resulting configuration will not be one that achieves the absolute minimum of the total energy but rather one that can’t be improved with local changes. Aaronson has carried out actual experiments (see Fig. 5.8), and saw that while this device often is successful for three or four pegs, it starts yielding suboptimal results once the number of pegs grows beyond that.

- **DNA computing.** People have suggested using the properties of DNA to do hard computational problems. The main advantage of DNA is the ability to potentially encode a lot of information in relatively small physical space, as well as compute on this information in a highly parallel manner. At the time of this writing, it was demonstrated that one can use DNA to store about \( 10^{16} \) bits of information in a region of radius about a millimeter, as opposed

*Figure 5.8: Scott Aaronson tests a candidate device for computing Steiner trees using soap bubbles.*
We were extremely conservative in the suggested parameters for the PECTT, having assumed that as many as $\ell^{-2} \cdot 10^{-6} \sim 10^6$ bits could potentially be stored in a millimeter radius region. This does not posit a real challenge to the PECTT but does suggest that one should be conservative about the choice of constant and not assume that current hard disk + silicon technologies are the absolute best possible.\footnote{We were extremely conservative in the suggested parameters for the PECTT, having assumed that as many as $\ell^{-2} \cdot 10^{-6} \sim 10^6$ bits could potentially be stored in a millimeter radius region.}

- **Continuous/real computers.** The physical world is often described using continuous quantities such as time and space, and people have suggested that analog devices might have direct access to computing with real-valued quantities and would be inherently more powerful than discrete models such as NAND machines. Whether the “true” physical world is continuous or discrete is an open question. In fact, we do not even know how to precisely phrase this question, let alone answer it. Yet, regardless of the answer, it seems clear that the effort to measure a continuous quantity grows with the level of accuracy desired, and so there is no “free lunch” or way to bypass the PECTT using such machines (see also this paper). Related to that are proposals known as “hypercomputing” or “Zeno’s computers” which attempt to use the continuity of time by doing the first operation in one second, the second one in half a second, the third operation in a quarter second and so on.. These fail for a similar reason to the one guaranteeing that Achilles will eventually catch the tortoise despite the original Zeno’s paradox.

- **Relativity computer and time travel.** The formulation above assumed the notion of time, but under the theory of relativity time is in the eye of the observer. One approach to solve hard problems is to leave the computer to run for a lot of time from his perspective, but to ensure that this is actually a short while from our perspective. One approach to do so is for the user to start the computer and then go for a quick jog at close to the speed of light before checking on its status. Depending on how fast one goes, few seconds from the point of view of the user might correspond to centuries in computer time (it might even finish updating its Windows operating system!). Of course the catch here is that the energy required from the user is proportional to how close one needs to get to the speed of light. A more interesting proposal is to use time travel via closed timelike curves (CTCs). In this case we could run an arbitrarily long computation by doing some calculations, remembering the current state, and the travelling back in time to continue where we left off. Indeed, if CTCs exist then we’d probably have to revise the PECTT (though in this case I will simply travel back in time and edit these notes, so I can claim I never conjectured it in the first place…)

- **Humans.** Another computing system that has been proposed as a counterexample to the PECTT is a 3 pound computer of about

...
0.1m radius, namely the human brain. Humans can walk around, talk, feel, and do others things that are not commonly done by NAND-CIRC programs, but can they compute partial functions that NAND-CIRC programs cannot? There are certainly computational tasks that at the moment humans do better than computers (e.g., play some video games, at the moment), but based on our current understanding of the brain, humans (or other animals) have no inherent computational advantage over computers. The brain has about \(10^{11}\) neurons, each operating in a speed of about 1000 operations per seconds. Hence a rough first approximation is that a Boolean circuit of about \(10^{14}\) gates could simulate one second of a brain’s activity.\(^8\) Note that the fact that such a circuit (likely) exists does not mean it is easy to find it. After all, constructing this circuit took evolution billions of years. Much of the recent efforts in artificial intelligence research is focused on finding programs that replicate some of the brain’s capabilities and they take massive computational effort to discover, these programs often turn out to be much smaller than the pessimistic estimates above. For example, at the time of this writing, Google’s neural network for machine translation has about \(10^4\) nodes (and can be simulated by a NAND-CIRC program of comparable size). Philosophers, priests and many others have since time immemorial argued that there is something about humans that cannot be captured by mechanical devices such as computers; whether or not that is the case, the evidence is thin that humans can perform computational tasks that are inherently impossible to achieve by computers of similar complexity.\(^9\)

• Quantum computation. The most compelling attack on the Physical Extended Church Turing Thesis comes from the notion of quantum computing. The idea was initiated by the observation that systems with strong quantum effects are very hard to simulate on a computer. Turning this observation on its head, people have proposed using such systems to perform computations that we do not know how to do otherwise. At the time of this writing, Scalable quantum computers have not yet been built, but it is a fascinating possibility, and one that does not seem to contradict any known law of nature. We will discuss quantum computing in much more detail in Chapter 22. Modeling quantum computation involves extending the model of Boolean circuits into Quantum circuits that have one more (very special) gate. However, the main take away is that while quantum computing does suggest we need to amend the PECTT, it does not require a complete revision of our worldview. Indeed, almost all of the content of this book remains the

\[3\] This is a very rough approximation that could be wrong to a few orders of magnitude in either direction. For one, there are other structures in the brain apart from neurons that one might need to simulate, hence requiring higher overhead. On the other hand, it is by no mean clear that we need to fully clone the brain in order to achieve the same computational tasks that it does.

\[9\] There are some well known scientists that have advocated that humans have inherent computational advantages over computers. See also this.
same regardless of whether the underlying computational model is Boolean circuits or quantum circuits.

**Remark 5.14 — Physical Extended Church-Turing Thesis and Cryptography.** While even the precise phrasing of the PECTT, let alone understanding its correctness, is still a subject of active research, some variants of it are already implicitly assumed in practice. Governments, companies, and individuals currently rely on cryptography to protect some of their most precious assets, including state secrets, control of weapon systems and critical infrastructure, securing commerce, and protecting the confidentiality of personal information.

In applied cryptography, one often encounters statements such as “cryptosystem $X$ provides 128 bits of security”. What such a statement really means that (a) it is conjectured that there is no Boolean circuit (or, equivalently, a NAND-CIRC program) of size much smaller than $2^{128}$ that can break $X$, and (b) we assume that no other physical mechanism can do better, and hence it would take roughly a $2^{128}$ amount of “resources” to break $X$. We say “conjectured” and not “proved” because, while we can phrase the statement that breaking the system cannot be done by an $s$-gate circuit as a precise mathematical conjecture, at the moment we are unable to prove such a statement for any non-trivial cryptosystem. This is related to the P vs NP question we will discuss in future chapters. We will explore Cryptography in Chapter 20.

**Lecture Recap**

- We can think of programs both as describing a process, as well as simply a list of symbols that can be considered as data that can be fed as input to other programs.
- We can write a NAND-CIRC program that evaluates arbitrary NAND-CIRC programs (or equivalently a circuit that evaluates other circuits). Moreover, the efficiency loss in doing so is not too large.
- We can even write a NAND-CIRC program that evaluates programs in other programming languages such as Python, C, Lisp, Java, Go, etc.
- By a leap of faith, we could hypothesize that the number of gates in the smallest circuit that computes a function $f$ captures roughly the amount of physical resources required to compute $f$. This statement is known as the Physical Extended Church-Turing Thesis (PECTT).
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- Boolean circuits (or equivalently AON-CIRC or NAND-CIRC programs) capture a surprisingly wide array of computational models. The strongest currently known challenge to the PECTT comes from the potential for using quantum mechanical effects to speed-up computation, a model known as quantum computers.

“What” (specification)

Function: \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \)

Algorithm/Program/Circuit:
- Boolean or NAND circuit \( C \),
- AON-CIRC or NAND-CIRC program \( P \)

\( 2^{O(s \log s)} \) circuits of size \( \leq s \)

“How” (implementation)

Every function computed by many circuits

Every program/circuit computes one function

Figure 5.9: A finite computational task is specified by a function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \). We can model a computational process using Boolean circuits (of varying gate sets) or straight-line program. Every function can be computed by many programs. We say that \( f \in \text{SIZE}_{n,m}(s) \) if there exists a NAND circuit of at most \( s \) gates (equivalently a NAND-CIRC program of at most \( s \) lines) that computes \( f \). Every function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) can be computed by a circuit of \( O(m \cdot 2^n / n) \) gates. Many functions such as multiplication, addition, solving linear equations, computing the shortest path in a graph, and others, can be computed by circuits of much fewer gates. In particular there is an \( O(s \log^2 s) \)-size circuit that computes the map \( C, x \mapsto C(x) \) where \( C \) is a string describing a circuit of \( s \) gates. However, the counting argument shows there do exist some functions \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) that require \( \Omega(m \cdot 2^n / n) \) gates to compute.

5.7 RECAP OF PART I: FINITE COMPUTATION

This chapter concludes the first part of this book that deals with finite computation (computing functions that map a fixed number of Boolean inputs to a fixed number of Boolean outputs). The main take-aways from Chapter 3, Chapter 4, and Chapter 5 are as follows (see also Fig. 5.9):

- We can formally define the notion of a function \( f : \{0, 1\}^n \rightarrow \{0, 1\}^m \) being computable using \( s \) basic operations. Whether these operations are AND/OR/NOT, NAND, or some other universal basis does not make much difference. We can describe such a computation either using a circuit or using a straight-line program.

- We define \( \text{SIZE}_{n,m}(s) \) to be the set of functions that are computable by NAND circuits of at most \( s \) gates. This set is equal to the set of functions computable by a NAND-CIRC program of at most \( s \) lines and pp to a constant factor in \( s \) (which we will not care about) this is also the same as the set of functions that are computable by a Boolean circuit of at most \( s \) AND/OR/NOT gates. The class \( \text{SIZE}_{n,m}(s) \) is a set of functions, not of programs/circuits.
• Every function \( f : \{0,1\}^n \rightarrow \{0,1\}^m \) can be computed using a circuit of at most \( O(m \cdot 2^n/n) \) gates. Some functions require at least \( \Omega(m \cdot 2^n/n) \) gates. We define \( \text{SIZE}_{n,m}(s) \) to be the set of functions from \( \{0,1\}^n \) to \( \{0,1\}^m \) that can be computed using at most \( s \) gates.

• We can describe a circuit/program \( P \) as a string. For every \( s \), there is a universal circuit/program \( U_s \) that can evaluate programs of length \( s \) given their description as strings. We can use this representation also to count the number of circuits of at most \( s \) gates and hence prove that some functions cannot be computed by circuit of smaller-than-exponential size.

• If there is a circuit of \( s \) gates that computes a function \( f \), then we can build a physical device to compute \( f \) using \( s \) basic components (such as transistors). The “Physical Extended Church-Turing Thesis” postulates postulates that the reverse direction is true as well: if \( f \) is a function for which every circuit requires at least \( s \) gates then that every physical device to compute \( f \) will require about \( s \) “physical resources”. The main challenge to the PECTT is quantum computing, which we will discuss in Chapter 22.

Sneak preview: In the next part we will discuss how to model computational tasks on unbounded inputs, which are specified using functions \( F : \{0,1\}^* \rightarrow \{0,1\}^* \) (or \( F : \{0,1\}^* \rightarrow \{0,1\} \)) that can take an unbounded number of Boolean inputs.

5.8 EXERCISES

Exercise 5.1 Which one of the following statements is false:

a. There is an \( O(s^3) \) line NAND-CIRC program that given as input program \( P \) of \( s \) lines in the list-of-tuples representation computes the output of \( P \) when all its input are equal to 1.

b. There is an \( O(s^3) \) line NAND-CIRC program that given as input program \( P \) of \( s \) characters encoded as a string of \( 7s \) bits using the ASCII encoding, computes the output of \( P \) when all its input are equal to 1.

c. There is an \( O(\sqrt{s}) \) line NAND-CIRC program that given as input program \( P \) of \( s \) lines in the list-of-tuples representation computes the output of \( P \) when all its input are equal to 1.

Exercise 5.2 — Equals function. For every \( k \in \mathbb{N} \), show that there is an \( O(k) \) line NAND-CIRC program that computes the function \( \text{EQUALS}_k : \{0,1\}^{2k} \rightarrow \{0,1\} \) where \( \text{EQUALS}(x,x') = 1 \) if and only if \( x = x' \)
Exercise 5.3 — Equal to constant function. For every $k \in \mathbb{N}$ and $x' \in \{0, 1\}^k$, show that there is an $O(k)$ line NAND-CIRC program that computes the function $EQL\{x'\} : \{0, 1\}^k \to \{0, 1\}$ that on input $x \in \{0, 1\}^k$ outputs 1 if and only if $x = x'$.

Exercise 5.4 — Counting lower bound for multibit functions. Prove that there exist a number $\delta > 0$ such that for every $n, m$ there exists a function $f : \{0, 1\}^n \to \{0, 1\}^m$ that requires at least $\delta m \cdot 2^n / n$ NAND gates to compute. See footnote for hint.\(^{10}\)

Exercise 5.5 — Size hierarchy theorem for multibit functions. Prove that there exists a number $C$ such that for every $n, m$ and $n + m < s < m \cdot 2^n / (Cn)$ there exists a function $f \in SIZE_{n,m}(C \cdot s) \setminus SIZE_{n,m}(s)$. See footnote for hint.\(^{11}\)

Exercise 5.6 — Efficient representation of circuits and a tighter counting upper bound. Use the ideas of Remark 5.4 to show that for every $\epsilon > 0$ and sufficiently large $s, n, m$,

$$|SIZE_{n,m}(s)| < 2^{(2+\epsilon)s \log s + n \log n + m \log s}.$$  \(5.11\)

Conclude that the implicit constant in Theorem 5.2 can be made arbitrarily close to 5. See footnote for hint.\(^{12}\)

Exercise 5.7 — Tighter counting lower bound. Prove that for every $\delta < 1/2$, if $n$ is sufficiently large then there exists a function $f : \{0, 1\}^n \to \{0, 1\}$ such that $f \notin SIZE_{n,1}(\frac{2^n}{n})$. See footnote for hint.\(^{13}\)

Exercise 5.8 — Random functions are hard. Suppose $n > 1000$ and that we choose a function $F : \{0, 1\}^n \to \{0, 1\}$ at random, choosing for every $x \in \{0, 1\}^n$ the value $F(x)$ to be the result of tossing an independent unbiased coin. Prove that the probability that there is a $2^n / (1000n)$ line program that computes $F$ is at most $2^{-100}$.\(^{14}\)

Exercise 5.9 The following is a tuple representing a NAND program:

$$(3, 1, ((3, 2, 2), (4, 1, 1), (5, 3, 4), (6, 2, 1), (7, 6, 6), (8, 0, 0), (9, 7, 8), (10, 5, 0), (11, 9, 10))).$$

1. Write a table with the eight values $P(000), P(001), P(010), P(011), P(100), P(101), P(110), P(111)$ in this order.

2. Describe what the programs does in words.
**Exercise 5.10 — EVAL with XOR.** For every sufficiently large \( n \), let \( E_n : \{0, 1\}^{n^2} \rightarrow \{0, 1\} \) be the function that takes an \( n^2 \)-length string that encodes a pair \((P, x)\) where \( x \in \{0, 1\}^n \) and \( P \) is a NAND program of \( n \) inputs, a single output, and at most \( n^{1.1} \) lines, and returns the output of \( P \) on \( x \).\(^{15}\) That is, \( E_n(P, x) = P(x) \).

Prove that for every sufficiently large \( n \), there does not exist an XOR circuit \( C \) that computes the function \( E_n \), where a XOR circuit has the XOR gate as well as the constants 0 and 1 (see Exercise 3.5). That is, prove that there is some constant \( n_0 \) such that for every \( n > n_0 \) and XOR circuit \( C \) of \( n^2 \) inputs and a single output, there exists a pair \((P, x)\) such that \( C(P, x) \neq E_n(P, x) \).

---

**Exercise 5.11 — Learning circuits (challenge, optional, assumes more background).**

(This exercise assumes background in probability theory and/or machine learning that you might not have at this point. Feel free to come back to it at a later point and in particular after going over Chapter 17.) In this exercise we will use our bound on the number of circuits of size \( s \) to show that (if we ignore the cost of computation) every such circuit can be learned from not too many training samples. Specifically, if we find a size-\( s \) circuit that classifies correctly a training set of \( O(s \log s) \) samples from some distribution \( D \), then it is guaranteed to do well on the whole distribution \( D \). Since Boolean circuits model very many physical processes (maybe even all of them, if the (controversial) physical extended Church-Turing thesis is true), this shows that all such processes could be learned as well (again, ignoring the computation cost of finding a classifier that does well on the training data).

Let \( D \) be any probability distribution over \( \{0, 1\}^n \) and let \( C \) be a NAND circuit with \( n \) inputs, one output, and size \( s \geq n \). Prove that there is some constant \( c \) such that with probability at least 0.999 the following holds: if \( m = cs \log s \) and \( x_0, \ldots, x_{m-1} \) are chosen independently from \( D \), then for every circuit \( C' \) such that \( C'(x_i) = C(x_i) \) on every \( i \in [m] \), \( \Pr_{x \sim D}[C'(x) \leq C(x)] \leq 0.99 \).

In other words, if \( C' \) is a so called “empirical risk minimizer” that agrees with \( C \) on all the training examples \( x_0, \ldots, x_{n-1} \), then it will also agree with \( C \) with high probability for samples drawn from the distribution \( D \) (i.e., it “generalizes”, to use Machine-Learning lingo). See footnote for hint.\(^{16}\)

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### 5.9 BIBLIOGRAPHICAL NOTES

The \( EVAL \) function is usually known as **universal circuit**. The implementation we describe is not the most efficient known. Valiant [Val76]...
first showed a universal circuit of $O(n \log n)$ size where $n$ is the size of
the input. Universal circuits have seen in recent years new motivations
due to their applications for cryptography, see [LMS16; GKS17].

While we’ve seen that “most” functions mapping $n$ bits to one bit
require circuits of exponential size $\Omega(2^{n/n})$, we actually do not know
of any explicit function for which we can prove that it requires, say,
at least $n^{100}$ or even $100n$ size. At the moment, strongest such lower
bound we know is that there are quite simple and explicit $n$-variable
functions that require at least $(5 - o(1))n$ lines to compute, see this
paper of Iwama et al as well as this more recent work of Kulikov et al.
Proving lower bounds for restricted models of circuits is an extremely
interesting research area, for which Jukna’s book [Juk12] (see also
Wegener [Weg87]) provides very good introduction and overview. I
learned of the proof of the size hierarchy theorem (Theorem 5.5) from
Sasha Golovnev.

Scott Aaronson’s blog post on how information is physical is a good
discussion on issues related to the physical extended Church-Turing
Physics. Aaronson’s survey on NP complete problems and physical
reality is also a great source for some of these issues, though might be
easier to read after we reach Chapter 14 on NP and NP-completeness.
II

UNIFORM COMPUTATION
6

Loops and infinity

“The algorithm is a finite answer to an infinite number of questions.”, Attributed to Stephen Kleene.

“The bounds of arithmetic were however outstepped the moment the idea of applying the [punched] cards had occurred; and the Analytical Engine does not occupy common ground with mere “calculating machines.””

... In enabling mechanism to combine together general symbols, in successions of unlimited variety and extent, a uniting link is established between the operations of matter and the abstract mental processes of the most abstract branch of mathematical science.”, Ada Augusta, countess of Lovelace, 1843

Learning Objectives:
- Learn the model of Turing machines, which can compute functions of arbitrary input lengths.
- See a programming-language description of Turing machines, using NAND-TM programs, which add loops and arrays to NAND-CIRC.
- See some basic syntactic sugar and equivalence of variants of Turing machines and NAND-TM programs.

The model of Boolean circuits (or equivalently, the NAND-CIRC programming language) has one very significant drawback: a Boolean circuit can only compute a finite function \( f \), and in particular since every gate has two inputs, a size \( s \) circuit can compute on an input of length at most \( 2^s \). This does not capture our intuitive notion of an algorithm as a single recipe to compute a potentially infinite function. For example, the standard elementary school multiplication algorithm is a single algorithm that multiplies numbers of all lengths, but yet we cannot express this algorithm as a single circuit, but rather need a different circuit (or equivalently, a NAND-CIRC program) for every input length (see Fig. 6.1).

Let us consider the case of the simple parity or XOR function XOR : \( \{0, 1\}^* \rightarrow \{0, 1\} \), where XOR\( (x) \) equals 1 iff the number of 1’s in \( x \) is odd. (In other words, \( \text{XOR}(x) = \sum_{i=0}^{\lfloor \log_2 |x| \rfloor} x_i \mod 2 \) for every \( x \in \{0, 1\}^* \).) As simple as it is, the XOR function cannot be computed by a NAND-CIRC program. Rather, for every \( n \), we can compute XOR\(_n\) (the restriction of XOR to \( \{0, 1\}^n \)) using a different NAND-CIRC program. For example, Fig. 6.2 presents the NAND-CIRC program (or equivalently the circuit) to compute XOR\(_5\).

Figure 6.1: Once you know how to multiply multi-digit numbers, you can do so for every number \( n \) of digits, but if you had to describe multiplication using NAND-CIRC programs or Boolean circuits, you would need a different program/circuit for every length \( n \) of the input.

Figure 6.2: The NAND circuit and NAND-CIRC program for computing the XOR of 5 bits. Note how the circuit for XOR\(_5\) merely repeats four times the circuit to compute the XOR of 2 bits.
This code for computing $XOR_5$ is rather repetitive, and more importantly, does not capture the fact that there is a single algorithm to compute the parity on all inputs. Typical programming language use the notion of loops to express such an algorithm, along the lines of:

```python
# s is the "running parity", initialized to 0
while i<len(X):
    u = NAND(s,X[i])
    v = NAND(s,u)
    w = NAND(X[i],u)
    s = NAND(v,w)
    i+= 1
Y[0] = s
```

Generally an algorithm is, as we quote above, “a finite answer to an infinite number of questions”. To express an algorithm we need to write down a finite set of instructions that will enable us to compute on arbitrarily long inputs. To describe and execute an algorithm we need the following components (see Fig. 6.3):

- The finite set of instructions to be performed.
- Some “local variables” or finite state used in the execution.
- A potentially unbounded working memory to store the input as well as any other values we may require later.
- While the memory is unbounded, at every single step we can only read and write to a finite part of it, and we need a way to address which are the parts we want to read from and write to.
- If we only have a finite set of instructions but our input can be arbitrarily long, we will need to repeat instructions (i.e., loop back). We need a mechanism to decide when we will loop and when we will halt.

In this chapter we will show how we can extend the model of Boolean circuits / straight-line programs so that it can capture these kinds of constructs. We will see two ways to do so:

- **Turing machines**, invented by Alan Turing in 1936, are an hypothetical abstract device that can yields a finite description of an algorithm that can handle arbitrarily long inputs.

- The **NAND-TM Programming language** extends NAND-CIRC with the notion of loops and arrays to obtain finite programs that can compute a function with arbitrarily long inputs.
It turns out that these two models are equivalent, and in fact they are equivalent to a great many other computational models including programming languages you may be familiar with such as C, Lisp, Python, JavaScript, etc. This notion, known as Turing equivalence or Turing completeness, will be discussed in Chapter 7. See Fig. 6.4 for an overview of the models presented in this chapter and Chapter 7.

**Remark 6.1 — Finite vs infinite computation.** Previously in this book we studied the computation of finite functions \( f : \{0, 1\}^n \to \{0, 1\}^m \). Such a function \( f \) can always be described by listing all the \( 2^n \) values it takes on inputs \( x \in \{0, 1\}^n \).

In this chapter we consider functions that take inputs of unbounded size, such as the function \( \text{XOR} : \{0, 1\}^* \to \{0, 1\} \) that maps \( x \) to \( \sum_{i=0}^{\left| x \right|-1} x_i \) mod 2. While we can describe \( \text{XOR} \) using a finite number of symbols (in fact we just did so in the previous sentence), it takes infinitely many possible inputs and so we cannot just write down all of its values. The same is true for many other functions capturing important computational tasks including addition, multiplication, sorting, finding paths in graphs, fitting curves to points, and so on and so forth.

To contrast with the finite case, we will sometimes call a function \( F : \{0, 1\}^* \to \{0, 1\} \) (or \( F : \{0, 1\}^* \to \{0, 1\}^* \) infinite but we emphasize that the functions we are interested in always take an input which is a finite string. It’s just that, unlike the finite case, this string can be arbitrarily long and is not fixed to some particular length \( n \).

Some texts present the task of computing a function \( F : \{0, 1\}^* \to \{0, 1\} \) as the task of deciding...
The “granddaddy” of all models of computation is the Turing Machine. Turing machines were defined in 1936 by Alan Turing in an attempt to formally capture all the functions that can be computed by human “computers” (see Fig. 6.6) that follow a well-defined set of rules, such as the standard algorithms for addition or multiplication.

Turing thought of such a person as having access to as much “scratch paper” as they need. For simplicity we can think of this scratch paper as a one dimensional piece of graph paper (or tape, as it is commonly referred to), which is divided to “cells”, where each “cell” can hold a single symbol (e.g., one digit or letter, and more generally some element of a finite alphabet). At any point in time, the person can read from and write to a single cell of the paper, and based on the contents can update his/her finite mental state, and/or move to the cell immediately to the left or right of the current one.

Turing modeled such a computation by a “machine” that maintains one of $k$ states. At each point in time the machine read from its “work tape” a single symbol from a finite alphabet $\Sigma$ and use that to update its state, write to tape, and possible move to an adjacent cell (see Fig. 6.9). To compute a function $F$ using this machine, we initialize the tape with the input $x \in \{0, 1\}^*$ and our goal is to ensure that the tape will contain the value $F(x)$ at the end of the computation. Specifically,
a computation of a Turing Machine $M$ with $k$ states and alphabet $\Sigma$ on input $x \in \{0,1\}^*$ proceeds as follows:

- Initially the machine is at state 0 (known as the “starting state”) and the tape is initialized to $\triangleright, x_0, ..., x_{n-1}, \emptyset, \emptyset, ...$. We use the symbol $\triangleright$ to denote the beginning of the tape, and the symbol $\emptyset$ to denote an empty cell. We will always assume that the alphabet $\Sigma$ is a (potentially strict) superset of $\{\triangleright, \emptyset, 0, 1\}$.
- The location $i$ to which the machine points to is set to 0.
- At each step, the machine reads the symbol $\sigma = T[i]$ that is in the $i^{th}$ location of the tape, and based on this symbol and its state $s$ decides on:
  - What symbol $\sigma'$ to write on the tape
  - Whether to move Left (i.e., $i \leftarrow i - 1$), Right (i.e., $i \leftarrow i + 1$), Stay in place, or Halt the computation.
  - What is going to be the new state $s \in [k]$
- The set of rules the Turing machine follows is known as its transition function.
- When the machine halts then its output is obtained by reading off the tape from the second location (just after the $\triangleright$) onwards, stopping at the first point where the symbol is not 0 or 1.

### 6.1.1 Extended example: A Turing machine for palindromes

Let $PAL$ (for palindromes) be the function that on input $x \in \{0,1\}^*$, outputs 1 if and only if $x$ is an (even length) palindrome, in the sense that $x = w_0 \cdots w_{n-1} w_{n-1} w_{n-2} \cdots w_0$ for some $n \in \mathbb{N}$ and $w \in \{0,1\}^n$.

We now show a Turing Machine $M$ that computes $PAL$. To specify $M$ we need to specify (i) $M$’s tape alphabet $\Sigma$ which should contain at least the symbols $0, 1, \triangleright, \emptyset, \times$, and (ii) $M$’s transition function which determines what action $M$ takes when it reads a given symbol while it is in a particular state.

In our case, $M$ will use the alphabet $\{0, 1, \triangleright, \emptyset, \times\}$ and will have $k = 14$ states. Though the states are simply numbers between 0 and $k - 1$, for convenience we will give them the following labels:

<table>
<thead>
<tr>
<th>State</th>
<th>Label</th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>START</td>
</tr>
<tr>
<td>1</td>
<td>RIGHT_0</td>
</tr>
<tr>
<td>2</td>
<td>RIGHT_1</td>
</tr>
<tr>
<td>3</td>
<td>LOOK_FOR_0</td>
</tr>
</tbody>
</table>

Figure 6.8: The components of a Turing Machine. Note how they correspond to the general components of algorithms as described in Fig. 6.3.
We describe the operation of our Turing Machine $M$ in words:

- $M$ starts in state START and will go right, looking for the first symbol that is 0 or 1. If we find $\varnothing$ before we hit such a symbol then we will move to the OUTPUT_1 state that we describe below.

- Once $M$ finds such a symbol $b \in \{0, 1\}$, $M$ deletes $b$ from the tape by writing the $\times$ symbol, it enters either the RIGHT_0 or RIGHT_1 mode according to the value of $b$ and starts moving rightwards until it hits the first $\varnothing$ or $\times$ symbol.

- Once we find this symbol we go into the state LOOK_FOR_0 or LOOK_FOR_1 depending on whether we were in the state RIGHT_0 or RIGHT_1 and make one left move.

- In the state LOOK_FOR_b, we check whether the value on the tape is $b$. If it is, then we delete it by changing its value to $\times$, and move to the state RETURN. Otherwise, we change to the OUTPUT_0 state.

- The RETURN state means we go back to the beginning. Specifically, we move leftward until we hit the first symbol that is not 0 or 1, in which case we change our state to START.

- The OUTPUT_b states mean that we are going to output the value $b$. In both these states we go left until we hit $\triangleright$. Once we do so, we make a right step, and change to the 1_AND_BLANK or 0_AND_BLANK states respectively. In the latter states, we write the corresponding value, and then move right and change to the BLANK_AND_STOP state, in which we write $\varnothing$ to the tape and halt.

The above description can be turned into a table describing for each one of the $13 \cdot 5$ combination of state and symbol, what the Turing machine will do when it is in that state and it reads that symbol. This table is known as the transition function of the Turing machine.
6.1.2 Turing machines: a formal definition

The formal definition of Turing machines is as follows:

**Definition 6.2 — Turing Machine.** A (one tape) Turing machine with \( k \) states and alphabet \( \Sigma \supseteq \{0, 1, \triangleright, \emptyset\} \) is represented by a transition function \( \delta_M : [k] \times \Sigma \rightarrow [k] \times \Sigma \times \{L, R, S, H\} \).

For every \( x \in \{0, 1\}^* \), the output of \( M \) on input \( x \), denoted by \( M(x) \), is the result of the following process:

- We initialize \( T \) to be the sequence \( \triangleright, x_0, x_1, \ldots, x_{n-1}, \emptyset, \emptyset, \ldots \) where \( n = |x| \). (That is, \( T[0] = \triangleright, T[i+1] = x_i \) for \( i \in [n] \), and \( T[i] = \emptyset \) for \( i > n \).)

- We also initialize \( i = 0 \) and \( s = 0 \).

- We then repeat the following process:

  1. Let \((s', \sigma', D) = \delta_M(s, T[i])\).
  2. Set \( s \rightarrow s' \), \( T[i] \rightarrow \sigma' \).
  3. If \( D = R \) then set \( i \rightarrow i + 1 \), if \( D = L \) then set \( i \rightarrow \max\{i - 1, 0\} \).
     (If \( D = S \) then we keep \( i \) the same.)
  4. If \( D = H \) then halt.

- The result of the process, which we denote by \( M(x) \), is the string \( T[1], \ldots, T[m] \) where \( m > 0 \) is the smallest integer such that \( T[m + 1] \notin \{0, 1\} \). If the process never ends then we write \( M(x) = \bot \).
You should make sure you see why this formal definition corresponds to our informal description of a Turing Machine. To get more intuition on Turing Machines, you can explore some of the online available simulators such as Martin Ugarte’s, Anthony Morphett’s, or Paul Rendell’s.

One should not confuse the transition function $\delta_M$ of a Turing machine $M$ with the function that the machine computes. The transition function $\delta_M$ is a \textit{finite} function, with $k|\Sigma|$ inputs and $4k|\Sigma|$ outputs. (Can you see why?) The machine can compute an \textit{infinite} function $F$ that takes as input a string $x \in \{0, 1\}^*$ of arbitrary length and might also produce an arbitrary length string as output.

In our formal definition, we identified the machine $M$ with its transition function $\delta_M$ since the transition function tells us everything we need to know about the Turing machine, and hence serves as a good mathematical representation of it. This choice of representation is somewhat arbitrary, and is based on our convention that the state space is always the numbers $\{0, \ldots, k-1\}$ with 0 as the starting state. Other texts use different conventions and so their mathematical definition of a Turing machine might look superficially different, but these definitions describe the same computational process and has the same computational powers. See Section 6.7 for a comparison between Definition 6.2 and the way Turing Machines are defined in texts such as Sipser [Sip97]. These definitions are equivalent despite their superficial differences.

### 6.1.3 Computable functions

We now turn to making one of the most important definitions in this book, that of \textit{computable functions}.

**Definition 6.3 — Computable functions.** Let $F : \{0, 1\}^* \to \{0, 1\}^*$ be a (total) function and let $M$ be a Turing machine. We say that $M$ \textit{computes} $F$ if for every $x \in \{0, 1\}^*$, $M(x) = F(x)$.

We say that a function $F$ is \textit{computable} if there exists a Turing machine $M$ that computes it.

Defining a function “computable” if and only if it can be computed by a Turing machine might seem “reckless” but, as we’ll see in Chapter 7, it turns out that being computable in the sense of Definition 6.3 is equivalent to being computable in essentially any reasonable model of computation. This is known as the \textit{Church Turing Thesis}. (Unlike the extended Church Turing Thesis which we discussed in Section 5.6,
the Church-Turing thesis itself is widely believed and there are no candidate devices that attack it.)

**Big Idea 8** We can precisely define what it means for a function to be computable by *any possible algorithm*.

This is a good point to remind the reader that *functions are not the same as programs*:

\[
\text{Functions } \neq \text{ Programs}. \tag{6.1}
\]

A Turing machine (or program) \( M \) can *compute* some function \( F \), but it is not the same as \( F \). In particular there can be more than one program to compute the same function. Being computable is a property of *functions*, not of machines.

We will often pay special attention to functions \( F : \{0, 1\}^* \to \{0, 1\} \) that have a single bit of output. Hence we give a special name for the set of functions of this form that are computable.

**Definition 6.4 — The class \( \mathbb{R} \).** We define \( \mathbb{R} \) be the set of all *computable* functions \( F : \{0, 1\}^* \to \{0, 1\} \).

**Remark 6.5 — Functions vs. languages.** Many texts use the terminology of "languages" rather than functions to refer to computational tasks. The name "language" has its roots in *formal language theory* as pursued by linguists such as Noam Chomsky. A *formal language* is a subset \( L \subseteq \{0, 1\}^* \) (or more generally \( L \subseteq \Sigma^* \) for some finite alphabet \( \Sigma \)). The *membership* or *decision* problem for a language \( L \), is the task of determining, given \( x \in \{0, 1\}^* \), whether or not \( x \in L \). A Turing machine \( M \) *decides* a language \( L \) if for every input \( x \in \{0, 1\}^* \), \( M(x) \) outputs 1 if and only if \( x \in L \). This is equivalent to computing the Boolean function \( F : \{0, 1\}^* \to \{0, 1\} \) defined as \( F(x) = 1 \) iff \( x \in L \). A language \( L \) is *decidable* if there is a Turing machine \( M \) that decides it. For historical reasons, some texts also call such a language *recursive* (which is the reason that the letter \( \mathbb{R} \) is often used to denote the set of computable Boolean functions / decidable languages defined in **Definition 6.4**).

In this book we stick to the terminology of *functions* rather than languages, but all definitions and results can be easily translated back and forth by using the equivalence between the function \( F : \{0, 1\}^* \to \{0, 1\} \) and the language \( L = \{x \in \{0, 1\}^* \mid F(x) = 1\} \).
6.1.4 Infinite loops and partial functions

One crucial difference between circuits/straight-line programs and Turing machines is the following. Looking at a NAND-CIRC program \( P \), we can always tell how many inputs and how many outputs it has (by simply looking at the \( X \) and \( Y \) variables). Furthermore, we are guaranteed that if we invoke \( P \) on any input then some output will be produced.

In contrast, given any Turing machine \( M \), we cannot determine a priori the length of the output. In fact, we don’t even know if an output would be produced at all! For example, it is very easy to come up with a Turing machine whose transition function never outsouts \( H \) and hence never halts.

If a machine \( M \) fails to stop and produce an output on some an input \( x \), then it cannot compute any total function \( F \), since clearly on input \( x \), \( M \) will fail to output \( F(x) \). However, \( P \) can still compute a partial function.\(^1\)

For example, consider the partial function \( \text{DIV} \) that on input a pair \((a, b)\) of natural numbers, outputs \([a/b]\) if \( b > 0 \), and is undefined otherwise. We can define a turing machine \( M \) that computes \( \text{DIV} \) on input \( a, b \) by outputting the first \( c = 0, 1, 2, \ldots \) such that \( cb \geq a \). If \( a > 0 \) and \( b = 0 \) then the machine \( M \) will never halt, but this is OK, since \( \text{DIV} \) is undefined on such inputs. If \( a = 0 \) and \( b = 0 \), the machine \( M \) will output 0, which is also OK, since we don’t care about what the program outputs on inputs on which \( \text{DIV} \) is undefined. Formally, we define computability of partial functions as follows:

\[\text{Definition 6.6 — Computable (partial or total) functions.} \] Let \( F \) be either a total or partial function mapping \( \{0, 1\}^* \) to \( \{0, 1\}^* \) and let \( M \) be a Turing machine. We say that \( M \) computes \( F \) if for every \( x \in \{0, 1\}^* \) on which \( F \) is defined, \( M(x) = F(x) \). We say that a (partial or total) function \( F \) is computable if there is a Turing machine that computes it.

Note that if \( F \) is a total function, then it is defined on every \( x \in \{0, 1\}^* \) and hence in this case, Definition 6.6 is identical to Definition 6.3.

\[\text{Remark 6.7 — Bot symbol.} \] We often use \( \bot \) as our special “failure symbol”. If a Turing machine \( M \) fails to halt on some input \( x \in \{0, 1\} \) then we denote this by \( M(x) = \bot \). This does not mean that \( M \) outputs some encoding of the symbol \( \bot \) but rather that \( M \) enters into an infinite loop when given \( x \) as input.

\(^1\) A partial function \( F \) from a set \( A \) to a set \( B \) is a function that is only defined on a subset of \( A \), (see Section 1.4.3). We can also think of such a function as mapping \( A \) to \( B \cup \{\bot\} \) where \( \bot \) is a special “failure” symbol such that \( F(a) = \bot \) indicates the function \( F \) is not defined on \( a \).
If a partial function $F$ is undefined on $x$ then can also write $F(x) = \bot$. Therefore one might think that Definition 6.6 can be simplified to requiring that $M(x) = F(x)$ for every $x \in \{0, 1\}$, which would imply that for every $x$, $M$ halts on $x$ if and only if $F$ is defined on $x$. However this is not the case: for a Turing Machine $M$ to compute a partial function $F$ it is not necessary for $M$ to enter an infinite loop on inputs $x$ on which $F$ is not defined. All that is needed is for $M$ to output $F(x)$ on $x$’s on which $F$ is defined: on other inputs it is OK for $M$ to output an arbitrary value such as 0, 1, or anything else, or not to halt at all. To borrow a term from the C programming language, on inputs $x$ on which $F$ is not defined, what $M$ does is “undefined behavior”.

6.2 TURING MACHINES AS PROGRAMMING LANGUAGES

The name “Turing machine”, with its “tape” and “head” evokes a physical object, while in contrast we think of a program as a piece of text. But we can think of a Turing machine as a program as well. For example, consider the Turing Machine $M$ of Section 6.1.1 that computes the function $PAL$ such that $PAL(x) = 1$ iff $x$ is a palindrome. We can also describe this machine as a program using the Python-like pseudocode of the form below

```
def PAL(Tape):
    head = 0
    state = 0 # START
    while (state != 12):
        if (state == 0 && Tape[head]=='0'):
            state = 3 # LOOK_FOR_0
            Tape[head] = 'x'
            head += 1 # move right
        if (state==0 && Tape[head]=='1')
            state = 4 # LOOK_FOR_1
            Tape[head] = 'x'
            head += 1 # move right
        ... # more if statements here
```

The particular details of this program are not important. What matters is that we can describe Turing machines as programs. Moreover, note that when translating a Turing machine into a program, the tape
becomes a list or array that can hold values from the finite set $\Sigma$. The head position can be thought of as an integer valued variable that can hold integers of unbounded size. The state is a local register that can hold one of a fixed number of values in $[k]$.

More generally we can think of every Turing Machine $M$ as equivalent to a program similar to the following:

```python
# Gets an array Tape initialized to # [">", x_0 , x_1 , ..., x_(n-1), " ", " ", ...]
def M(Tape):
    state = 0
    i = 0 # holds head location
    while (True):
        # Move head, modify state, write to tape
        # based on current state and cell at head
        # below are just examples for how program looks
        # for a particular transition function
        if Tape[i]=="0" and state==7:  #
            δ_M(7,"0")=(19,"1","R")
            i += 1
            Tape[i]="1"
            state = 19
        elif Tape[i]==">" and state == 13:  #
            δ_M(13,">")=(15,"0","S")
            Tape[i]="0"
            state = 15
        elif ...
        elif Tape[i]==">" and state == 29:  #
            δ_M(29,">")=(.,.,"H")
            break # Halt
```

If we wanted to use only Boolean (i.e., 0/1-valued) variables then we can encode the state variables using $\lceil \log k \rceil$ bits. Similarly, we can represent each element of the alphabet $\Sigma$ using $\ell = \lceil \log |\Sigma| \rceil$ bits and hence we can replace the $\Sigma$-valued array $\text{Tape}[]$ with $\ell$ Boolean-valued arrays $\text{Tape0}[]$, $\ldots$, $\text{Tape}\ell[]$.

### 6.2.1 The NAND-TM Programming language

We now introduce the NAND-TM programming language, which aims to capture the power of a Turing machine in a programming language formalism. Just like the difference between Boolean circuits and Turing Machines, the main difference between NAND-TM and NAND-CIRC is that NAND-TM models a single uniform algorithm that can
compute a function that takes inputs of arbitrary lengths. To do so, we extend the NAND-CIRC programming language with two constructs:

- **Loops:** NAND-CIRC is a straight-line programming language—a NAND-CIRC program of \( s \) lines takes exactly \( s \) steps of computation and hence in particular cannot even touch more than \( 3s \) variables. **Loops** allow us to capture in a short program the instructions for a computation that can take an arbitrary amount of time.

- **Arrays:** A NAND-CIRC program of \( s \) lines touches at most \( 3s \) variables. While we can use variables with names such as `Foo_17` or `Bar[22]`, they are not true arrays, since the number in the identifier is a constant that is “hardwired” into the program.

Thus a good way to remember NAND-TM is using the following informal equation:

\[
\text{NAND-TM} = \text{NAND-CIRC} + \text{loops} + \text{arrays} \quad (6.2)
\]

**Remark 6.8** — NAND-CIRC + loops + arrays = everything. As we will see, adding loops and arrays to NAND-CIRC is enough to capture the full power of all programming languages! Hence we could replace “NAND-TM” with any of Python, C, Javascript, OCaml, etc. in the lefthand side of (6.2). But we’re getting ahead of ourselves: this issue will be discussed in Chapter 7.

Concretely, the NAND-TM programming language adds the following features on top of NAND-CIRC (see Fig. 6.10):

- We add a special integer valued variable \( i \). All other variables in NAND-TM are Boolean valued (as in NAND-CIRC).
• Apart from \( i \) NAND-TM has two kinds of variables: *scalars* and *arrays*. Scalar variables hold one bit (just as in NAND-CIRC). Array variables hold an unbounded number of bits. At any point in the computation we can access the array variables at the location indexed by \( i \) using \( \text{Foo}[i] \). We cannot access the arrays at locations other the one pointed to by \( i \).

• We use the convention that arrays always start with a capital letter, and scalar variables (which are never indexed with \( i \)) start with lowercase letters. Hence \( \text{Foo} \) is an array and \( \text{bar} \) is a scalar variable.

• The input and output \( X \) and \( Y \) are now considered arrays with values of zeroes and ones. (There are also two other special arrays \( X_{\text{nonblank}} \) and \( Y_{\text{nonblank}} \), see below.)

• We add a special \textbf{MODANDJUMP} instruction that takes two boolean variables \( a, b \) as input and does the following:

  - If \( a = 1 \) and \( b = 1 \) then \textbf{MODANDJUMP}(\( a, b \)) increments \( i \) by one and jumps to the first line of the program.
  - If \( a = 0 \) and \( b = 1 \) then \textbf{MODANDJUMP}(\( a, b \)) decrements \( i \) by one and jumps to the first line of the program. (If \( i \) is already equal to 0 then it stays at 0.)
  - If \( a = 1 \) and \( b = 0 \) then \textbf{MODANDJUMP}(\( a, b \)) jumps to the first line of the program without modifying \( i \).
  - If \( a = b = 0 \) then \textbf{MODANDJUMP}(\( a, b \)) halts execution of the program.

• The \textbf{MODANDJUMP} instruction always appears in the last line of a NAND-TM program and nowhere else.

**Default values.** We need one more convention to handle “default values”. Turing machines have the special symbol \( \varnothing \) to indicate that tape location is “blank” or “uninitialized”. In NAND-TM there is no such symbol, and all variables are Boolean, containing either 0 or 1. All variables and locations of arrays are default to 0 if they have not been initialized to another value. To keep track of whether a 0 in an array corresponds to a true zero or to an uninitialized cell, a programmer can always add to an array \( \text{Foo} \) a “companion array” \( \text{Foo}_{\text{nonblank}} \) and set \( \text{Foo}_{\text{nonblank}}[i] \) to 1 whenever the \( i \)’th location is initialized. In particular we will use this convention for the input and output arrays \( X \) and \( Y \). A NAND-TM program has four special arrays \( X \), \( X_{\text{nonblank}} \), \( Y \), and \( Y_{\text{nonblank}} \). When a NAND-TM program is executed on input \( x \in \{0, 1\}^* \) of length \( n \), the first \( n \) cells of the array \( X \) are initialized to \( x_0, \ldots, x_{n-1} \), and the first \( n \) cells of the array \( X_{\text{nonblank}} \) are initialized to 1. (All uninitialized cells default to 0.) The output of
a NAND-TM program is the string \( Y[0], \ldots, Y[m-1] \) where \( m \) is the smallest integer such that \( Y_{nonblank}[m] = 0 \). A NAND-TM program gets called with \( X \) and \( X_{nonblank} \) initialized to contain the input, and writes to \( Y \) and \( Y_{nonblank} \) to produce the output.

Formally, NAND-TM programs are defined as follows:

**Definition 6.9 — NAND-TM programs.** A NAND-TM program consists of a sequence of lines of the form \( \text{foo} = \text{NAND} (\text{bar}, \text{blah}) \) ending with a line of the form \( \text{MODANDJMP} (\text{foo}, \text{bar}) \), where \( \text{foo}, \text{bar}, \text{blah} \) are either scalar variables (sequences of letters, digits, and underscores) or array variables of the form \( \text{Foo}[i] \) (starting with capital letter and indexed by \( i \)). The program has the array variables \( X, X_{nonblank}, Y, Y_{nonblank} \) and the index variable \( i \) built in, and can use additional array and scalar variables.

If \( P \) is a NAND-TM program and \( x \in \{0, 1\}^* \) is an input then an execution of \( P \) on \( x \) is the following process:

1. The arrays \( X \) and \( X_{nonblank} \) are initialized by \( X[i] = x_i \) and \( X_{nonblank}[i] = 1 \) for all \( i \in |x| \). All other variables and cells are initialized to 0. The index variable \( i \) is also initialized to 0.

2. The program is executed line by line, when the last line \( \text{MODANDJMP} (\text{foo}, \text{bar}) \) is executed then we do as follows:
   
   a. If \( \text{foo} = 1 \) and \( \text{bar} = 0 \) then jump to the first line without modifying the value of \( i \).
   
   b. If \( \text{foo} = 1 \) and \( \text{bar} = 1 \) then increment \( i \) by one and jump to the first line.
   
   c. If \( \text{foo} = 0 \) and \( \text{bar} = 1 \) then decrement \( i \) by one (unless it is already zero) and jump to the first line.
   
   d. If \( \text{foo} = 0 \) and \( \text{bar} = 0 \) then halt and output \( Y[0], \ldots, Y[m-1] \) where \( m \) is the smallest integer such that \( Y_{nonblank}[m] = 0 \).

### 6.2.2 Sneak peak: NAND-TM vs Turing machines

As the name implies, NAND-TM programs are a direct implementation of Turing machines in programming language form. We will show the equivalence below but you can already see how the components of Turing machines and NAND-TM programs correspond to one another:
### Table 6.2: Turing Machine and NAND-TM analogs

<table>
<thead>
<tr>
<th>Turing Machine</th>
<th>NAND-TM program</th>
</tr>
</thead>
<tbody>
<tr>
<td><strong>State:</strong> single register that takes values in $[k]$</td>
<td><strong>Scalar variables:</strong> Several variables such as foo, bar etc. each taking values in ${0, 1}$.</td>
</tr>
<tr>
<td><strong>Tape:</strong> One tape containing values in a finite set $\Sigma$. Potentially infinite but $T[t]$ defaults to $\emptyset$ for all locations $t$ that have not been accessed.</td>
<td><strong>Arrays:</strong> Several arrays such as $\text{Foo}$, $\text{Bar}$ etc. for each such array $\text{Arr}$ and index $j$, the value of $\text{Arr}$ at position $j$ is either 0 or 1. The value defaults to 0 for position that have not been written to.</td>
</tr>
<tr>
<td><strong>Head location:</strong> A number $i \in \mathbb{N}$ that encodes the position of the head.</td>
<td><strong>Index variable:</strong> The variable $i$ that can be used to access the arrays.</td>
</tr>
<tr>
<td><strong>Accessing memory:</strong> At every step the Turing machine has access to its local state, but can only access the tape at the position of the current head location.</td>
<td><strong>Accessing memory:</strong> At every step a NAND-TM program has access to all the scalar variables, but can only access the arrays at the location $i$ of the index variable.</td>
</tr>
<tr>
<td><strong>Control of location:</strong> In each step the machine can move the head location by at most one position.</td>
<td><strong>Control of index variable:</strong> In each iteration of its main loop the program can modify the index $i$ by at most one.</td>
</tr>
</tbody>
</table>

### 6.2.3 Examples

We now present some examples of NAND-TM programs

**Example 6.10 — XOR in NAND-TM.** The following is a NAND-TM program to compute the XOR function on inputs of arbitrary length. That is $\text{XOR} : \{0, 1\}^* \rightarrow \{0, 1\}$ such that $\text{XOR}(x) = \sum_{i=0}^{\lfloor |x| / 2 \rfloor} x_i \mod 2$ for every $x \in \{0, 1\}^*$.

```plaintext
temp_0 = \text{NAND}(X[0], X[0])
Y_nonblank[0] = \text{NAND}(X[0], temp_0)
temp_2 = \text{NAND}(X[i], Y[0])
temp_3 = \text{NAND}(X[i], temp_2)
temp_4 = \text{NAND}(Y[0], temp_2)
Y[0] = \text{NAND}(temp_3, temp_4)
\text{MODANDJUMP}(X\_nonblank[i], X\_nonblank[i])
```
Example 6.11 — Increment in NAND-TM. We now present NAND-TM program to compute the increment function. That is, \( \text{INC} : \{0, 1\}^* \rightarrow \{0, 1\}^* \) such that for every \( x \in \{0, 1\}^n \), \( \text{INC}(x) \) is the \( n + 1 \) bit long string \( y \) such that if \( X = \sum_{i=0}^{n-1} x_i \cdot 2^i \) is the number represented by \( x \), then \( y \) is the (least-significant digit first) binary representation of the number \( X + 1 \).

We start by showing the program using the “syntactic sugar” we’ve seen before of using shorthand for some NAND-CIRC programs we have seen before to compute simple functions such as IF, XOR and AND (as well as the constant one function as well as the function COPY that just maps a bit to itself).

```plaintext
carry = IF(started, carry, one(started))
started = one(started)
Y[i] = XOR(X[i], carry)
carry = AND(X[i], carry)
Y_nonblank[i] = one(started)
MODANDJUMP(X_nonblank[i], X_nonblank[i])
```

The above is not, strictly speaking, a valid NAND-TM program. If we “open up” all of the syntactic sugar, we get the following valid program to compute this syntactic sugar.

```plaintext
temp_0 = NAND(started, started)
temp_1 = NAND(started, temp_0)
temp_2 = NAND(started, started)
temp_3 = NAND(temp_1, temp_2)
temp_4 = NAND(carry, started)
carry = NAND(temp_3, temp_4)
temp_6 = NAND(started, started)
started = NAND(started, temp_6)
temp_8 = NAND(X[i], carry)
temp_9 = NAND(X[i], temp_8)
temp_10 = NAND(carry, temp_8)
Y[i] = NAND(temp_9, temp_10)
temp_12 = NAND(X[i], carry)
carry = NAND(temp_12, temp_12)
temp_14 = NAND(started, started)
Y_nonblank[i] = NAND(started, temp_14)
MODANDJUMP(X_nonblank[i], X_nonblank[i])
```
Working out the above two example can go a long way towards understanding the NAND-TM language. See the appendix and our GitHub repository for a full specification of the NAND-TM language.

6.3 EQUIVALENCE OF TURING MACHINES AND NAND-TM PROGRAMS

Given the above discussion, it might not be surprising that Turing machines turn out to be equivalent to NAND-TM programs. Indeed, we designed the NAND-TM language to have this property. Nevertheless, this is an important result, and the first of many other such equivalence results we will see in this book.

**Theorem 6.12** — Turing machines and NAND-TM programs are equivalent. For every $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, $F$ is computable by a NAND-TM program $P$ if and only if there is a Turing Machine $M$ that computes $F$.

**Proof Idea:**

To prove such an equivalence theorem, we need to show two directions. We need to be able to (1) transform a Turing machine $M$ to a NAND-TM program $P$ that computes the same function as $P$ and (2) transform a NAND-TM program $P$ into a Turing machine $M$ that computes the same function as $P$.

The idea of the proof is illustrated in Fig. 6.11. To show (1), given a Turing machine $M$, we will create a NAND-TM program $P$ that will have an array Tape for the tape of $M$ and scalar (i.e., non array) variable(s) state for the state of $M$. Specifically, since the state of a Turing machine is not in $\{0, 1\}$ but rather in a larger set $[k]$, we will use $\lceil \log k \rceil$ variables state_0, ..., state_{\lceil \log k \rceil} – 1 variables to store the representation of the state. Similarly, to encode the larger alphabet $\Sigma$ of the tape, we will use $\lceil \log |\Sigma| \rceil$ arrays Tape_0, ..., Tape_{\lceil \log |\Sigma| \rceil} – 1, such that the $i^{th}$ location of these arrays encodes the $i^{th}$ symbol in the tape for every tape. Using the fact that every function can be computed by a NAND-CIRC program, we will be able to compute the transition function of $M$, replacing moving left and right by decrementing and incrementing $i$ respectively.

We show (2) using very similar ideas. Given a program $P$ that uses $a$ array variables and $b$ scalar variables, we will create a Turing machine with about $2^b$ states to encode the values of scalar variables, and an alphabet of about $2^a$ so we can encode the arrays using our tape. (The reason the sizes are only "about" $2^a$ and $2^b$ is that we will need to add some symbols and steps for bookkeeping purposes.)
Turing Machine $M$ will simulate each iteration of the program $P$ by updating its state and tape accordingly.

* Figure 6.11: Comparing a Turing Machine to a NAND-TM program. Both have an unbounded memory component (the tape for a Turing machine, and the arrays for a NAND-TM program), as well as a constant local memory (state for a Turing machine, and scalar variables for a NAND-TM program). Both can only access at each step one location of the unbounded memory, this is the “head” location for a Turing machine, and the value of the index variable $i$ for a NAND-TM program.

Proof of Theorem 6.12. We start by proving the “if” direction of Theorem 6.12. Namely we show that given a Turing machine $M$, we can find a NAND-TM program $P_M$ such that for every input $x$, if $M$ halts on input $x$ with output $y$ then $P_M(x) = y$. Since our goal is just to show such a program $P_M$ exists, we don’t need to write out the full code of $P_M$ line by line, and can take advantage of our various “syntactic sugar” in describing it.

The key observation is that by Theorem 4.12 we can compute every finite function using a NAND-CIRC program. In particular, consider the transition function $\delta_M : [k] \times \Sigma \rightarrow [k] \times \Sigma \times \{L, R, S, H\}$ of our Turing Machine. We can encode the its components as follows:

- We encode $[k]$ using $\{0, 1\}^\ell$ and $\Sigma$ using $\{0, 1\}^{\ell'}$, where $\ell = \lceil \log k \rceil$ and $\ell' = \lceil \log |\Sigma| \rceil$.

- We encode the set $\{L, R, S, H\}$ using $\{0, 1\}^2$. We will choose the encode $L \mapsto 01$, $R \mapsto 11$, $S \mapsto 10$, $H \mapsto 00$. (This conveniently corresponds to the semantics of the MODANDJUMP operation.)

Hence we can identify $\delta_M$ with a function $\overline{M} : \{0, 1\}^{\ell + \ell'} \rightarrow \{0, 1\}^{\ell + \ell' + 2}$, mapping strings of length $\ell + \ell'$ to strings of length $\ell + \ell' + 2$. By Theorem 4.12 there exists a finite length NAND-CIRC program $\text{ComputeM}$ that computes this function $\overline{M}$. The NAND-TM program to simulate $M$ will essentially be the following:
Algorithm 6.13 — NAND-TM program to simulate TM $M$.

**Input:** $x \in \{0, 1\}^*$

**Output:** $M(x)$ if $M$ halts on $x$. Otherwise go into infinite loop.

1. # We use variables state_0 ... state_{ℓ − 1} to encode $M$’s state
2. # We use arrays Tape_0[] ... Tape_{ℓ'} − 1[] to encode $M$’s tape
3. # We omit the initial and final "book keeping" to copy

**Input:** to Tape and copy

**Output:** from Tape

4. # Use the fact that transition is finite and computable by NAND-CIRC program:
5. state_0 ... state_{ℓ − 1}, Tape_0[i]... Tape_{ℓ'} − 1[i], dir0,dir1 ← TRANSITION( state_0 ... state_{ℓ − 1}, Tape_0[i]... Tape_{ℓ'} − 1[i], dir0,dir1 )
6. MODANDJMP(dir0,dir1)

Every step of the main loop of the above program perfectly mimics the computation of the Turing Machine $M$ and so the program carries out exactly the definition of computation by a Turing Machine as per Definition 6.2.

For the other direction, suppose that $P$ is a NAND-TM program with $s$ lines, $\ell$ scalar variables, and $\ell'$ array variables. We will show that there exists a Turing machine $M_P$ with $2^\ell + C$ states and alphabet $\Sigma$ of size $C' + 2\ell'$ that computes the same functions as $P$ (where $C, C'$ are some constants to be determined later).

Specifically, consider the function $\overline{P} : \{0, 1\}^\ell \times \{0, 1\}^{\ell'} \rightarrow \{0, 1\}^\ell \times \{0, 1\}^{\ell'}$ that on input the contents of $P$’s scalar variables and the contents of the array variables at location $i$ in the beginning of an iteration, outputs all the new values of these variables at the last line of the iteration, right before the MODANDJUMP instruction is executed.

If foo and bar are the two variables that are used as input to the MODANDJUMP instruction, then this means that based on the values of these variables we can compute whether $i$ will increase, decrease or stay the same, and whether the program will halt or jump back to the beginning. Hence a Turing machine can simulate an execution of $P$ in one iteration using a finite function applied to its alphabet. The overall operation of the Turing machine will be as follows:

1. The machine $M_P$ encodes the contents of the array variables of $P$ in its tape, and the contents of the scalar variables in (part of) its state. Specifically, if $P$ has $\ell$ local variables and $t$ arrays, then the state space of $M$ will be large enough to encode all $2^\ell$ assignments
to the local variables and the alphabet $\Sigma$ of $M$ will be large enough to encode all $2^t$ assignments for the array variables at each location. The head location corresponds to the index variable $i$.

2. Recall that every line of the program $P$ corresponds to reading and writing either a scalar variable, or an array variable at the location $i$. In one iteration of $P$ the value of $i$ remains fixed, and so the machine $M$ can simulate this iteration by reading the values of all array variables at $i$ (which are encoded by the single symbol in the alphabet $\Sigma$ located at the $i$-th cell of the tape), reading the values of all scalar variables (which are encoded by the state), and updating both. The transition function of $M$ can output $L$, $S$, $R$ depending on whether the values given to the $\text{MODANDJMP}$ operation are $01$, $10$ or $11$ respectively.

3. When the program halts (i.e., $\text{MODANDJMP}$ gets $00$) then the Turing machine will enter into a special loop to copy the results of the $Y$ array into the output and then halt. We can achieve this by adding a few more states.

The above is not a full formal description of a Turing Machine, but our goal is just to show that such a machine exists. One can see that $M_P$ simulates every step of $P$, and hence computes the same function as $P$.

\[\text{R \ Remark 6.14 — Running time equivalence (optional). If we examine the proof of Theorem 6.12 then we can see that the every iteration of the loop of a NAND-TM program corresponds to one step in the execution of the Turing machine. We will come back to this question of measuring number of computation steps later in this course. For now the main take away point is that NAND-TM programs and Turing Machines are essentially equivalent in power even when taking running time into account.}\]

### 6.3.1 Specification vs implementation (again)

Once you understand the definitions of both NAND-TM programs and Turing Machines, Theorem 6.12 is fairly straightforward. Indeed, NAND-TM programs are not as much a different model from Turing Machines as they are simply a reformulation of the same model using programming language notation. You can think of the difference between a Turing machine and a NAND-TM program as the difference between representing a number using decimal or binary notation. In
contrast, the difference between a function $F$ and a Turing machine that computes $F$ is much more profound: it is like the difference between the equation $x^2 + x = 12$ and the number 3 that is a solution for this equation. For this reason, while we take special care in distinguishing functions from programs or machines, we will often identify the two latter concepts. We will move freely between describing an algorithm as a Turing machine or as a NAND-TM program (as well as some of the other equivalent computational models we will see in Chapter 7 and beyond).

<table>
<thead>
<tr>
<th>Setting</th>
<th>Specification</th>
<th>Implementation</th>
</tr>
</thead>
<tbody>
<tr>
<td>Finite computation</td>
<td>Functions mapping ${0, 1}^n$ to ${0, 1}^m$</td>
<td>Circuits, Straightline programs</td>
</tr>
<tr>
<td>Infinite computation</td>
<td>Functions mapping ${0, 1}^<em>$ to ${0, 1}$ or to ${0, 1}^</em>$.</td>
<td>Algorithms, Turing Machines, Programs</td>
</tr>
</tbody>
</table>

### 6.4 NAND-TM SYNTACTIC SUGAR

Just like we did with NAND-CIRC in Chapter 4, we can use "syntactic sugar" to make NAND-TM programs easier to write. For starters, we can use all of the syntactic sugar of NAND-CIRC, and so have access to macro definitions and conditionals (i.e., if/then). But we can go beyond this and achieve for example:

- Inner loops such as the while and for operations common to many programming languages
- Multiple index variables (e.g., not just $i$ but we can add $j$, $k$, etc.).
- Arrays with more than one dimension (e.g., $\text{Foo}[i][j]$, $\text{Bar}[i][j][k]$ etc.)

In all of these cases (and many others) we can implement the new feature as mere "syntactic sugar" on top of standard NAND-TM, which means that the set of functions computable by NAND-TM with this feature is the same as the set of functions computable by standard NAND-TM. Similarly, we can show that the set of functions computable by Turing Machines that have more than one tape, or tapes of more dimensions than one, is the same as the set of functions computable by standard Turing machines.
6.4.1 “GOTO” and inner loops

We can implement more advanced looping constructs than the simple MODANDJUMP. For example, we can implement GOTO. A GOTO statement corresponds to jumping to a certain line in the execution. For example, if we have code of the form

"start": do foo
    GOTO("end")
"skip": do bar
"end": do blah

then the program will only do foo and blah as when it reaches the line GOTO("end") it will jump to the line labeled with "end". We can achieve the effect of GOTO in NAND-TM using conditionals. In the code below, we assume that we have a variable pc that can take strings of some constant length. This can be encoded using a finite number of Boolean variables \( pc_0, pc_1, \ldots, pc_{k-1} \), and so when we write below \( pc = "label" \) what we mean is something like \( pc_0 = 0, pc_1 = 1, \ldots \) (where the bits 0, 1, ... correspond to the encoding of the finite string "label" as a string of length \( k \)). We also assume that we have access to conditional (i.e., if statements), which we can emulate using syntactic sugar in the same way as we did in NAND-CIRC.

To emulate a GOTO statement, we will first modify a program P of the form

do foo
do bar
do blah

...to have the following form (using syntactic sugar for if):

\[
\begin{align*}
\text{pc} & = \"line1\" \\
\text{if} \ (\text{pc}==\"line1\"): & \quad \text{do} \ \text{foo} \\
& \quad \quad \text{pc} = \"line2\" \\
\text{if} \ (\text{pc}==\"line2\"): & \quad \text{do} \ \text{bar} \\
& \quad \quad \text{pc} = \"line3\" \\
\text{if} \ (\text{pc}==\"line3\"): & \quad \text{do} \ \text{blah}
\end{align*}
\]

These two programs do the same thing. The variable pc corresponds to the “program counter” and tells the program which line to execute next. We can see that if we wanted to emulate a GOTO("line3") then we could simply modify the instruction \( \text{pc} = \"line2\" \) to be \( \text{pc} = \"line3\" \).
In NAND-CIRC we could only have GOTOs that go forward in the code, but since in NAND-TM everything is encompassed within a large outer loop, we can use the same ideas to implement GOTO’s that can go backwards, as well as conditional loops.

**Other loops.** Once we have GOTO, we can emulate all the standard loop constructs such as while, do ... until or for in NAND-TM as well. For example, we can replace the code

```plaintext
while foo:
    do blah
do bar
```

with

```plaintext
"loop":
    if NOT(foo): GOTO("next")
    do blah
    GOTO("loop")
"next":
    do bar
```

---

**Remark 6.15 — GOTO’s in programming languages.** The GOTO statement was a staple of most early programming languages, but has largely fallen out of favor and is not included in many modern languages such as Python, Java, Javascript. In 1968, Edsger Dijkstra wrote a famous letter titled “Go to statement considered harmful.” (see also Fig. 6.12). The main trouble with GOTO is that it makes analysis of programs more difficult by making it harder to argue about invariants of the program.

When a program contains a loop of the form:

```plaintext
for j in range(100):
    do something
    do blah
```

you know that the line of code do blah can only be reached if the loop ended, in which case you know that j is equal to 100, and might also be able to argue other properties of the state of the program. In contrast, if the program might jump to do blah from any other point in the code, then it’s very hard for you as the programmer to know what you can rely upon in this code. As Dijkstra said, such invariants are important because “our intellectual powers are rather geared to master static relations and .. our powers to visualize processes evolving in time are relatively poorly developed.”
and so “we should ... do ... our utmost best to shorten the conceptual gap between the static program and the dynamic process.”

That said, GOTO is still a major part of lower level languages where it is used to implement higher level looping constructs such as while and for loops. For example, even though Java doesn’t have a GOTO statement, the Java Bytecode (which is a lower level representation of Java) does have such a statement. Similarly, Python bytecode has instructions such as POP_JUMP_IF_TRUE that implement the GOTO functionality, and similar instructions are included in many assembly languages. The way we use GOTO to implement a higher level functionality in NAND-TM is reminiscent of the way these various jump instructions are used to implement higher level looping constructs.

6.5 UNIFORMITY, AND NAND VS NAND-TM (DISCUSSION)

While NAND-TM adds extra operations over NAND-CIRC, it is not exactly accurate to say that NAND-TM programs or Turing machines are “more powerful” than NAND-CIRC programs or Boolean circuits. NAND-CIRC programs, having no loops, are simply not applicable for computing functions with an unbounded number of inputs. Thus, to compute a function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$ using NAND-CIRC (or equivalently, Boolean circuits) we need a collection of programs/circuits: one for every input length.

The key difference between NAND-CIRC and NAND-TM is that NAND-TM allows us to express the fact that the algorithm for computing parities of length-100 strings is really the same one as the algorithm for computing parities of length-5 strings (or similarly the fact that the algorithm for adding $n$-bit numbers is the same for every $n$, etc.). That is, one can think of the NAND-TM program for general parity as the “seed” out of which we can grow NAND-CIRC programs for length 10, length 100, or length 1000 parities as needed.

This notion of a single algorithm that can compute functions of all input lengths is known as uniformity of computation and hence we think of Turing machines / NAND-TM as uniform model of computation, as opposed to Boolean circuits or NAND-CIRC which is a nonuniform model, where we have to specify a different program for every input length.

Looking ahead, we will see that this uniformity leads to another crucial difference between Turing machines and circuits. Turing machines can have inputs and outputs that are longer than the description of the machine as a string and in particular there exists a Turing machine that can “self replicate” in the sense that it can print its own
code. This notion of “self replication”, and the related notion of “self reference” is crucial to many aspects of computation, as well of course to life itself, whether in the form of digital or biological programs.

For now, what you ought to remember is the following differences between uniform and non uniform computational models:

- **Non uniform computational models**: Examples are NAND-CIRC programs and Boolean circuits. These are models where each individual program/circuit can compute a finite function $f : \{0, 1\}^n \rightarrow \{0, 1\}^m$. We have seen that every finite function can be computed by some program/circuit. To discuss computation of an infinite function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$ we need to allow a sequence $\{P_n\}_{n \in \mathbb{N}}$ of programs/circuits (one for every input length), but this does not capture the notion of a single algorithm to compute the function $F$.

- **Uniform computational models**: Examples are Turing machines and NAND-TM programs. These are model where a single program/machine can take inputs of arbitrary length and hence compute an infinite function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$. The number of steps that a program/machine takes on some input is not a priori bounded in advance and in particular there is a chance that it will enter into an infinite loop. Unlike the nonuniform case, we have not shown that every infinite function can be computed by some NAND-TM program/Turing Machine. We will come back to this point in Chapter 8.

---

**Lecture Recap**

- Turing machines capture the notion of a single algorithm that can evaluate functions of every input length.
- They are equivalent to NAND-TM programs, which add loops and arrays to NAND-CIRC.
- Unlike NAND-CIRC or Boolean circuits, the number of steps that a Turing machine takes on a given input is not fixed in advance. In fact, a Turing machine or a NAND-TM program can enter into an infinite loop on certain inputs, and not halt at all.

---

### 6.6 EXERCISES

**Exercise 6.1 — Explicit NAND TM programming.** Produce the code of a (syntactic-sugar free) NAND-TM program $P$ that computes the (unbounded input length) Majority function $Maj : \{0, 1\}^* \rightarrow \{0, 1\}$ where for every $x \in \{0, 1\}^*$, $Maj(x) = 1$ if and only if $\sum_{i=0}^{\left\lfloor \frac{|x|}{2} \right\rfloor} x_i > |x|/2$. We say “produce” rather than “write” because you do not have to write
the code of $P$ by hand, but rather can use the programming language of your choice to compute this code.

**Exercise 6.2 — Computable functions examples.** Prove that the following functions are computable. For all of these functions, you do not have to fully specify the Turing Machine or the NAND-TM program that computes the function, but rather only prove that such a machine or program exists:

1. $\text{INC} : \{0, 1\}^* \rightarrow \{0, 1\}$ which takes as input a representation of a natural number $n$ and outputs the representation of $n + 1$.

2. $\text{ADD} : \{0, 1\}^* \rightarrow \{0, 1\}$ which takes as input a representation of a pair of natural numbers $(n, m)$ and outputs the representation of $n + m$.

3. $\text{MULT} : \{0, 1\}^* \rightarrow \{0, 1\}^*$, which takes a representation of a pair of natural numbers $(n, m)$ and outputs the representation of $nm$.

4. $\text{SORT} : \{0, 1\}^* \rightarrow \{0, 1\}^*$ which takes as input the representation of a list of natural numbers $(a_0, \ldots, a_{n-1})$ and returns its sorted version $(b_0, \ldots, b_{n-1})$ such that for every $i \in [n]$ there is some $j \in [n]$ with $b_i = a_j$ and $b_0 \leq b_1 \leq \cdots \leq b_{n-1}$.

**Exercise 6.3 — Two index NAND-TM.** Define NAND-TM’ to be the variant of NAND-TM where there are two index variables $i$ and $j$. Arrays can be indexed by either $i$ or $j$. The operation $\text{MODANDJMP}$ takes four variables $a, b, c, d$ and uses the values of $c, d$ to decide whether to increment $j$, decrement $j$ or keep it in the same value (corresponding to $01$, $10$, and $00$ respectively). Prove that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, $F$ is computable by a NAND-TM program if and only if $F$ is computable by a NAND-TM’ program.

**Exercise 6.4 — Two tape Turing machines.** Define a two tape Turing machine to be a Turing machine which has two separate tapes and two separate heads. At every step, the transition function gets as input the location of the cells in the two tapes, and can decide whether to move each head independently. Prove that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, $F$ is computable by a standard Turing Machine if and only if $F$ is computable by a two-tape Turing machine.

**Exercise 6.5 — Two dimensional arrays.** Define NAND-TM” to be the variant of NAND-TM where just like NAND-TM’ defined in Exercise 6.3
there are two index variables $i$ and $j$, but now the arrays are *two dimensional* and so we index an array $\text{Foo}$ by $\text{Foo}[i][j]$. Prove that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, $F$ is computable by a NAND-TM program if and only if $F$ is computable by a NAND-TM” program.

**Exercise 6.6 — Two tape Turing machines.** Define a *two-dimensional Turing machine* to be a Turing machine in which the tape is *two dimensional*. At every step the machine can move Up, Down, Left, Right, or Stay. Prove that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, $F$ is computable by a standard Turing Machine if and only if $F$ is computable by a two-dimensional Turing machine.

**Exercise 6.7** Prove the following closure properties of the set $R$ defined in Definition 6.4:

1. If $F \in R$ then the function $G(x) = 1 - F(x)$ is in $R$.
2. If $F, G \in R$ then the function $H(x) = F(x) \lor G(x)$ is in $R$.
3. If $F \in R$ then the function $F^*$ in in $R$ where $F^*$ is defined as follows: $F^*(x) = 1$ iff there exist some strings $w_0, \ldots, w_{k-1}$ such that $x = w_0w_1\cdots w_{k-1}$ and $F(w_i) = 1$ for every $i \in [k]$.
4. If $F \in R$ then the function
   \[
   G(x) = \begin{cases} 
   1 & \exists y \in \{0, 1\}^n \text{ such that } F(xy) = 1 \\
   0 & \text{otherwise} \end{cases} \tag{6.3}
   
   is in $R$.

**Exercise 6.8 — Oblivious Turing Machines (challenging).** Define a Turing Machine $M$ to be *oblivious* if its head movement are independent of its input. That is, we say that $M$ is oblivious if there existe an infinite sequence $MOVE \in \{L, R, S\}^\infty$ such that for every $x \in \{0, 1\}^*$, the movements of $M$ when given input $x$ (up until the point it halts, if such point exists) are given by $MOVE_0, MOVE_1, MOVE_2, \ldots$.

Prove that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, if $F$ is computable then it is computable by an oblivious Turing machine. See footnote for hint.\(^3\)

**Exercise 6.9 — Single vs multiple bit.** Prove that for every $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$, the function $F$ is computable if and only if the following func-
tion $G : \{0, 1\}^* \to \{0, 1\}$ is computable, where $G$ is defined as follows:

$$G(x, i, \sigma) = \begin{cases} F(x)_i & i < |F(x)|, \sigma = 0 \\ 1 & i < |F(x)|, \sigma = 1 \\ 0 & i \geq |F(x)| \end{cases}$$

**Exercise 6.10 — Uncomputability via counting.** Recall that $R$ is the set of all total functions from $\{0, 1\}^*$ to $\{0, 1\}$ that are computable by a Turing machine (see Definition 6.4). Prove that $R$ is countable. That is, prove that there exists a one-to-one map $D_1N : R \to \mathbb{N}$. You can use the equivalence between Turing machines and NAND-TM programs.

**Exercise 6.11 — Not every function is computable.** Prove that the set of all total functions from $\{0, 1\}^* \to \{0, 1\}$ is not countable. You can use the results of Section 2.3.1. (We will see an explicit uncomputable function in Chapter 8.)

### 6.7 BIBLIOGRAPHICAL NOTES

Augusta Ada Byron, countess of Lovelace (1815-1852) lived a short but turbulent life, though is today most well known for her collaboration with Charles Babbage (see [Ste87] for a biography). Ada took an immense interest in Babbage’s *analytical engine*, which we mentioned in Chapter 3. In 1842-3, she translated from Italian a paper of Menabrea on the engine, adding copious notes (longer than the paper itself). The quote in the chapter’s beginning is taken from Nota A in this text. Lovelace’s notes contain several examples of programs for the analytical engine, and because of this she has been called “the world’s first computer programmer” though it is not clear whether they were written by Lovelace or Babbage himself [Hol01]. Regardless, Ada was clearly one of very few people (perhaps the only one outside of Babbage himself) to fully appreciate how significant and revolutionary the idea of mechanizing computation truly is.

The books of Shetterly [She16] and Sobel [Sob17] discuss the history of human computers (who were female, more often than not) and their important contributions to scientific discoveries in astronomy and space exploration.

Alan Turing was one of the intellectual giants of the 20th century. He was not only the first person to define the notion of computation, but also invented and used some of the world’s earliest computational devices as part of the effort to break the *Enigma* cipher during World War II, saving millions of lives. Tragically, Turing committed suicide in 1954, following his conviction in 1952 for homosexual acts and a court-mandated hormonal treatment. In 2009, British prime minister
Gordon Brown made an official public apology to Turing, and in 2013 Queen Elizabeth II granted Turing a posthumous pardon. Turing’s life is the subject of a great book and a mediocre movie.

Sipser’s text [Sip97] defines a Turing machine is as a seven tuple consisting of the state space, input alphabet, tape alphabet, transition function, starting state, accepting state, and rejecting state. Superficially this looks like a very different definition than Definition 6.2 but it is simply a different representation of the same concept, just as a graph can be represented in either adjacency list or adjacency matrix form.

One difference is that Sipser considers a general set of states $Q$ that is not necessarily of the form $Q = \{0, 1, 2, \ldots, k - 1\}$ for some natural number $k > 0$. Sipser also restricts his attention to Turing machines that output only a single bit and therefore designates two special halting states: the “0 halting state” (often known as the rejecting state) and the other as the “1 halting state” (often known as the accepting state). Thus instead of writing 0 or 1 on an output tape, the machine will enter into one of these states and halt. This again makes no difference to the computational power, though we prefer to consider the more general model of multi-bit outputs. (Sipser presents the basic task of a Turing machine as that of deciding a language as opposed to computing a function, but these are equivalent, see Remark 6.5.)

Sipser considers also functions with input in $\Sigma^*$ for an arbitrary alphabet $\Sigma$ (and hence distinguishes between the input alphabet which he denotes as $\Sigma$ and the tape alphabet which he denotes as $\Gamma$), while we restrict attention to functions with binary strings as input. Again this is not a major issue, since we can always encode an element of $\Sigma$ using a binary string of length $\log \lceil |\Sigma| \rceil$. Finally (and this is a very minor point) Sipser requires the machine to either move left or right in every step, without the Stay operation, though staying in place is very easy to emulate by simply moving right and then back left.

Another definition used in the literature is that a Turing machine $M$ recognizes a language $L$ if for every $x \in L$, $M(x) = 1$ and for every $x \notin L$, $M(x) \in \{0, \bot\}$. A language $L$ is recursively enumerable if there exists a Turing machine $M$ that recognizes it, and the set of all recursively enumerable languages is often denoted by RE. We will not use this terminology in this book.

One of the first programming-language formulations of Turing machines was given by Wang [Wan57]. Our formulation of NAND-TM is aimed at making the connection with circuits more direct, with the eventual goal of using it for the Cook-Levin Theorem, as well as results such as $P \subseteq P/poly$ and $\text{BPP} \subseteq P/poly$. The website esolangs.org features a large variety of esoteric Turing-complete programming languages. One of the most famous of them is Brainf*ck.
7
Equivalent models of computation

“All problems in computer science can be solved by another level of indirection”, attributed to David Wheeler.

“Because we shall later compute with expressions for functions, we need a distinction between functions and forms and a notation for expressing this distinction. This distinction and a notation for describing it, from which we deviate trivially, is given by Church.”, John McCarthy, 1960 (in paper describing the LISP programming language)

So far we have defined the notion of computing a function using Turing machines, which are not a close match to the way computation is done in practice. In this chapter we justify this choice by showing that the definition of computable functions will remain the same under a wide variety of computational models. This notion is known as Turing completeness or Turing equivalence and is one of the most fundamental facts of computer science. In fact, a widely believed claim known as the Church-Turing Thesis holds that every “reasonable” definition of computable function is equivalent to being computable by a Turing machine. We discuss the Church-Turing Thesis and the potential definitions of “reasonable” in Section 7.8.

Some of the main computational models we discuss in this chapter include:

- **RAM Machines**: Turing Machines do not correspond to standard computing architectures that have Random Access Memory (RAM). The mathematical model of RAM machines is much closer to actual computers, but we will see that it is equivalent in power to Turing Machines. We also discuss a programming language variant of RAM machines, which we call NAND-RAM. The equivalence of Turing Machines and RAM machines enables demonstrating the


Turing Equivalence of many popular programming languages, including all general-purpose languages used in practice such as C, Python, JavaScript, etc.

- **Cellular Automata:** Many natural and artificial systems can be modeled as collections of simple components, each evolving according to simple rules based on its state and the state of its immediate neighbors. One well-known such example is Conway’s Game of Life. To prove that cellular automata are equivalent to Turing machines we introduce the tool of configurations of Turing Machines. These have other applications, and in particular are used in Chapter 10 to prove Gödel’s Incompleteness Theorem: a central result in mathematics.

- **λ calculus:** The λ calculus is a model for expressing computation that originates from the 1930’s, though it is closely connected to functional programming languages widely used today. Showing the equivalence of λ calculus to Turing Machines involves a beautiful technique to eliminate recursion known as the “Y Combinator”.

![Some Turing-equivalent models. All of these are equivalent in power to Turing Machines (or equivalently NAND-TM programs) in the sense that they can compute exactly the same class of functions. All of these are models for computing infinite functions that take inputs of unbounded length. In contrast, Boolean circuits / NAND-CIRC programs can only compute finite functions and hence are not Turing complete.](image)

### 7.1 RAM MACHINES AND NAND-RAM

One of the limitations of Turing Machines (and NAND-TM programs) is that we can only access one location of our arrays/tape at a time. If the head is at position 22 in the tape and we want to access the 957-th position then it will take us at least 923 steps to get there. In contrast, almost every programming language has a formalism for directly accessing memory locations. Actual physical computers also provide so called Random Access Memory (RAM) which can be thought of as a large array of Memory, such that given an index \( p \) (i.e., memory address, or a pointer), we can read from and write to the \( p \)th location of Memory. (“Random access memory” is quite a misnomer since it has nothing to do with probability, but since it is a standard term in both the theory and practice of computing, we will use it as well.)
The computational model that models access to such a memory is the RAM machine (sometimes also known as the Word RAM model), as depicted in Fig. 7.2. The memory of a RAM machine is an array of unbounded size where each cell can store a single word, which we think of as a string in \( \{0,1\}^w \) and also (equivalently) as a number in \( [2^w] \). For example, many modern computing architectures use 64 bit words, in which every memory location holds a string in \( \{0,1\}^{64} \) which can also be thought of as a number between 0 and \( 2^{64} - 1 = 9,223,372,036,854,775,807 \). The parameter \( w \) is known as the word size. In practice often \( w \) is a fixed number such as 64, but when doing theory we model \( w \) as a parameter that can depend on the input length or number of steps. (You can think of \( 2^w \) as roughly corresponding to the largest memory address that we use in the computation.) In addition to the memory array, a RAM machine also contains a constant number of registers \( r_0, \ldots, r_{k-1} \), each of which can also contain a single word.

The operations a RAM machine can carry out include:

- **Data movement**: Load data from a certain cell in memory into a register or store the contents of a register into a certain cell of memory. RAM machine can directly access any cell of memory without having to move the “head” (as Turing machines do) to that location. That is, in one step a RAM machine can load into register \( r_i \) the contents of the memory cell indexed by register \( r_j \), or store into the memory cell indexed by register \( r_j \) the contents of register \( r_i \).

- **Computation**: RAM machines can carry out computation on registers such as arithmetic operations, logical operations, and comparisons.

- **Control flow**: As in the case of Turing machines, the choice of what instruction to perform next can depend on the state of the RAM machine, which is captured by the contents of its register.

We will not give a formal definition of RAM Machines, though the bibliographical notes section (Section 7.10) contains sources for such definitions. Just as the NAND-TM programming language models Turing machines, we can also define a NAND-RAM programming language that models Turing machines. The NAND-RAM programming language extends NAND-TM by adding the following features:

- The variables of NAND-RAM are allowed to be (non negative) integer valued rather than only Boolean as is the case in NAND-TM. That is, a scalar variable \( \text{foo} \) holds an non negative integer in \( \mathbb{N} \) (rather than only a bit in \( \{0,1\} \)), and an array variable \( \text{Bar} \) holds...
an array of integers. As in the case of RAM machines, we will not allow integers of unbounded size. Concretely, each variable holds a number between 0 and \( T - 1 \), where \( T \) is the number of steps that have been executed by the program so far. (You can ignore this restriction for now: if we want to hold larger numbers, we can simply execute dummy instructions; it will be useful in later chapters.)

- We allow indexed access to arrays. If \( \text{foo} \) is a scalar and \( \text{Bar} \) is an array, then \( \text{Bar}[\text{foo}] \) refers to the location of \( \text{Bar} \) indexed by the value of \( \text{foo} \). (Note that this means we don’t need to have a special index variable \( i \) any more.)

- As is often the case in programming languages, we will assume that for Boolean operations such as \( \text{NAND} \), a zero valued integer is considered as false, and a nonzero valued integer is considered as true.

- In addition to \( \text{NAND} \), NAND-RAM also includes all the basic arithmetic operations of addition, subtraction, multiplication, (integer) division, as well as comparisons (equal, greater than, less than, etc.).

- NAND-RAM includes conditional statements if/then as part of the language.

- NAND-TM contains looping constructs such as goto, while and do as part of the language.

A full description of the NAND-RAM programming language is in the appendix. However, the most important fact you need to know about NAND-RAM is that you actually don’t need to know much about NAND-RAM at all, since it is equivalent in power to Turing machines:

**Theorem 7.1** — Turing Machines (aka NAND-TM programs) and RAM machines (aka NAND-RAM programs) are equivalent. For every function \( F : \{0, 1\}^* \rightarrow \{0, 1\}^* \), \( F \) is computable by a NAND-TM program if and only if \( F \) is computable by a NAND-RAM program.

Since NAND-TM programs are equivalent to Turing machines, and NAND-RAM programs are equivalent to RAM machines, Theorem 7.1 shows that all these four models are equivalent to one another.

**Proof Idea:**

Clearly NAND-RAM is only more powerful than NAND-TM, and so if a function \( F \) is computable by a NAND-TM program then it can be computed by a NAND-RAM program. The challenging direction is...
to transform a NAND-RAM program $P$ to an equivalent NAND-TM program $Q$. To describe the proof in full we will need to cover the full formal specification of the NAND-RAM language, and show how we can implement every one of its features as syntactic sugar on top of NAND-TM.

This can be done but going over all the operations in detail is rather tedious. Hence we will focus on describing the main ideas behind this transformation. (See also Fig. 7.3.) NAND-RAM generalizes NAND-TM in two main ways: (a) adding indexed access to the arrays (i.e., $\text{Foo}[\text{bar}]$ syntax) and (b) moving from Boolean valued variables to integer valued ones. The transformation has two steps:

1. **Indexed access of bit arrays**: We start by showing how to handle (a). Namely, we show how we can implement in NAND-TM the operation $\text{Setindex}()$ such that if $\text{Bar}$ is an array that encodes some integer $j$, then after executing $\text{Setindex}()$ the value of $i$ will equal to $j$. This will allow us to simulate syntax of the form $\text{Foo}[\text{Bar}]$ by $\text{Setindex}()$ followed by $\text{Foo}[i]$.

2. **Two dimensional bit arrays**: We then show how we can use “syntactic sugar” to augment NAND-TM with two dimensional arrays. That is, have two indices $i$ and $j$ and two dimensional arrays, such that we can use the syntax $\text{Foo}[i][j]$ to access the $(i,j)$-th location of $\text{Foo}$.

3. **Arrays of integers**: Finally we will encode a one dimensional array $\text{Arr}$ of integers by a two dimensional $\text{Arrbin}$ of bits. The idea is simple: if $a_{i,0}, \ldots, a_{i,\ell}$ is a binary (prefix-free) representation of $\text{Arr}[i]$, then $\text{Arrbin}[i][j]$ will be equal to $a_{i,j}$.

Once we have arrays of integers, we can use our usual syntactic sugar for functions, $\text{GOTO}$ etc. to implement the arithmetic and control flow operations of NAND-RAM.

\[ \star \]

**Remark 7.2 — RAM machines / NAND-RAM and assembly language (optional).** RAM machines correspond quite closely to actual microprocessors such as those in the Intel x86 series that also contains a large primary memory and a constant number of small registers. This is of course no accident: RAM machines aim at modeling more closely than Turing machines the architecture of actual computing systems, which largely follows the so called von Neumann architecture as described in the report [Neu45]. As a result, NAND-RAM is similar in its general outline to assembly languages such as x86 or NIPS. These assembly languages all have instructions to (1) move data from registers to mem-
ory, (2) perform arithmetic or logical computations on registers, and (3) conditional execution and loops ("if" and "goto", commonly known as "branches" and "jumps" in the context of assembly languages).

The main difference between RAM machines and actual microprocessors (and correspondingly between NAND-RAM and assembly languages) is that actual microprocessors have a fixed word size \( w \) so that all registers and memory cells hold numbers in \( [2^w] \) (or equivalently strings in \( \{0, 1\}^w \)). This number \( w \) can vary among different processors, but common values are either 32 or 64. As a theoretical model, RAM machines do not have this limitation, but we rather let \( w \) be the logarithm of our running time (which roughly corresponds to its value in practice as well). Actual microprocessors also have a fixed number of registers (e.g., 14 general purpose registers in x86-64) but this does not make a big difference with RAM machines. It can be shown that RAM machines with as few as two registers are as powerful as full-fledged RAM machines that have an arbitrarily large constant number of registers.

Of course actual microprocessors have many features not shared with RAM machines as well, including parallelism, memory hierarchies, and many others. However, RAM machines do capture actual computers to a first approximation and so (as we will see), the running time of an algorithm on a RAM machine (e.g., \( O(n) \) vs \( O(n^2) \)) is strongly correlated with its practical efficiency.

### 7.2 THE GORY DETAILS (OPTIONAL)

We do not show the full formal proof of Theorem 7.1 but focus on the most important parts: implementing indexed access, and simulating two dimensional arrays with one dimensional ones. Even these are already quite tedious to describe, as will not be surprising to anyone that has ever written a compiler. Hence you can feel free to merely skim this section. The important point is not for you to know all details by heart but to be convinced that in principle it is possible to transform a NAND-RAM program to an equivalent NAND-TM program, and even be convinced that, with sufficient time and effort, you could do it if you wanted to.

#### 7.2.1 Indexed access in NAND-TM

In NAND-TM we can only access our arrays in the position of the index variable \( i \), while NAND-RAM has integer-valued variables and can use them for indexed access to arrays, of the form \( \text{Foo}[\text{bar}] \). To implement indexed access in NAND-TM, we will encode integers in our
arrays using some prefix-free representation (see Section 2.4.2), and then have a procedure Setindex(Bar) that sets i to the value encoded by Bar. We can simulate the effect of Foo[Bar] using Setindex(Bar) followed by Foo[i].

Implementing Setindex(Bar) can be achieved as follows:

1. We initialize an array Azero such that Azero[0] = 1 and Azero[j] = 0 for all j > 0. (This can be easily done in NAND-TM as all uninitialized variables default to zero.)

2. Set i to zero, by decrementing it until we reach the point where Azero[i] = 1.

3. Let Temp be an array encoding the number 0.

4. We use GOTO to simulate an inner loop of the form: while Temp ≠ Bar, increment Temp.

5. At the end of the loop, i is equal to the value encoded by Bar.

In NAND-TM code (using some syntactic sugar), we can implement the above operations as follows:

```nand-tm
# assume Azero is an array such that Azero[0]=1
# and Azero[j]=0 for all j>0

# set i to 0.
LABEL("zero_idx")
dir0 = zero
dir1 = one
# corresponds to i <- i-1
GOTO("zero_idx",NOT(Azero[i]))
...
# zero out temp
# (code below assumes a specific prefix-free encoding in which 10 is the "end marker")
Temp[0] = 1
Temp[1] = 0
# set i to Bar, assume we know how to increment, compare
LABEL("increment_temp")
cond = EQUAL(Temp,Bar)
dir0 = one
dir1 = one
# corresponds to i <- i+1
INC(Temp)
GOTO("increment_temp",cond)
# if we reach this point, i is number encoded by Bar
```
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...  

# final instruction of program
MODANDJUMP(dir0,dirl)

7.2.2 Two dimensional arrays in NAND-TM

To implement two dimensional arrays, we want to embed them in a one dimensional array. The idea is that we come up with a *one to one* function $\text{embed} : \mathbb{N} \times \mathbb{N} \rightarrow \mathbb{N}$, and so embed the location $(i, j)$ of the two dimensional array Two in the location $\text{embed}(i, j)$ of the array One.

Since the set $\mathbb{N} \times \mathbb{N}$ seems “much bigger” than the set $\mathbb{N}$, a priori it might not be clear that such a one to one mapping exists. However, once you think about it more, it is not that hard to construct. For example, you could ask a child to use scissors and glue to transform a 10” by 10” piece of paper into a 1” by 100” strip. This is essentially a one to one map from $[10] \times [10]$ to $[100]$. We can generalize this to obtain a one to one map from $[n] \times [n]$ to $[n^2]$ and more generally a one to one map from $\mathbb{N} \times \mathbb{N}$ to $\mathbb{N}$. Specifically, the following map $\text{embed}$ would do (see Fig. 7.4):

$$
\text{embed}(x, y) = \frac{1}{2}(x + y)(x + y + 1) + x .
$$

(7.1)

Exercise 7.2 asks you to prove that $\text{embed}$ is indeed one to one, as well as computable by a NAND-TM program. (The latter can be done by simply following the grade-school algorithms for multiplication, addition, and division.) This means that we can replace code of the form Two[Foo][Bar] = something (i.e., access the two dimensional array Two at the integers encoded by the one dimensional arrays Foo and Bar) by code of the form:

Blah = embed(Foo,Bar)
Setindex(Blah)
Two[i] = something

7.2.3 All the rest

Once we have two dimensional arrays and indexed access, simulating NAND-RAM with NAND-TM is just a matter of implementing the standard algorithms for arithmetic operations and comparisons in NAND-TM. While this is cumbersome, it is not difficult, and the end result is to show that every NAND-RAM program $P$ can be simulated by an equivalent NAND-TM program $Q$, thus completing the proof of Theorem 7.1.

**Remark 7.3 — Recursion in NAND-RAM (advanced).** One concept that appears in many programming languages but we did not include in NAND-RAM programs is...
**7.3 TURING EQUIVALENCE (DISCUSSION)**

Any of the standard programming languages such as C, Java, Python, Pascal, Fortran have very similar operations to NAND-RAM. (Indeed, ultimately they can all be executed by machines which have a fixed number of registers and a large memory array.) Hence using Theorem 7.1, we can simulate any program in such a programming language by a NAND-TM program. In the other direction, it is a fairly easy programming exercise to write an interpreter for NAND-TM in any of the above programming languages. Hence we can also simulate NAND-TM programs (and so by Theorem 6.12, Turing machines) using these programming languages. This property of being equivalent in power to Turing Machines / NAND-TM is called Turing Equivalent (or sometimes Turing Complete). Thus all programming languages we are familiar with are Turing equivalent.\(^1\)

---

\(^1\) Some programming languages have fixed (even if extremely large) bounds on the amount of memory they can access, which formally prevent them from being applicable to computing infinite functions and hence simulating Turing machines. We ignore such issues in this discussion and assume access to some storage device without a fixed upper bound on its capacity.
7.3.1 The “Best of both worlds” paradigm

The equivalence between Turing Machines and RAM machines allows us to choose the most convenient language for the task at hand:

- When we want to prove a theorem about all programs/algorithms, we can use Turing machines (or NAND-TM) since they are simpler and easier to analyze. In particular, if we want to show that a certain function cannot be computed, then we will use Turing machines.

- When we want to show that a function can be computed we can use RAM machines or NAND-RAM, because they are easier to program in and correspond more closely to high level programming languages we are used to. In fact, we will often describe NAND-RAM programs in an informal manner, trusting that the reader can fill in the details and translate the high level description to the precise program. (This is just like the way people typically use informal or “pseudocode” descriptions of algorithms, trusting that their audience will know to translate these descriptions to code if needed.)

Our usage of Turing Machines / NAND-TM and RAM Machines / NAND-RAM is very similar to the way people use in practice high and low level programming languages. When one wants to produce a device that executes programs, it is convenient to do so for very simple and “low level” programming language. When one wants to describe an algorithm, it is convenient to use as high level a formalism as possible.

Big Idea 9 Using equivalence results such as those between Turing and RAM machines, we can “have our cake and eat it too”. We can use a simpler model such as Turing machines when we want to prove something cannot be done, and use a feature-rich model such as RAM machines when we want to prove something can be done.

7.3.2 Let’s talk about abstractions.

“The programmer is in the unique position that ... he has to be able to think in terms of conceptual hierarchies that are much deeper than a single mind ever needed to face before.”, Edsger Dijkstra, “On the cruelty of really teaching computing science”, 1988.

At some point in any theory of computation course, the instructor and students need to have the talk. That is, we need to discuss the level
of abstraction in describing algorithms. In algorithms courses, one typically describes algorithms in English, assuming readers can “fill in the details” and would be able to convert such an algorithm into an implementation if needed. For example, Algorithm 7.4 is a high level description of the breadth first search algorithm.

Algorithm 7.4 — Breadth First Search.

**Input:** Graph $G$, vertices $u$, $v$

**Output:** “connected” when $u$ is connected to $v$ in $G$, “disconnected”

1: Initialize empty queue $Q$.
2: Put $u$ in $Q$
3: while $Q$ is not empty do
4: Remove top vertex $w$ from $Q$
5: if $w = v$ then
6: return “connected”
7: end if
8: Mark $w$
9: Add all unmarked neighbors of $w$ to $Q$.
10: end while
11: return “disconnected”

If we wanted to give more details on how to implement breadth first search in a programming language such as Python or C (or NAND-RAM / NAND-TM for that matter), we would describe how we implement the queue data structure using an array, and similarly how we would use arrays mark vertices. We call such an “intermediate level” description an implementation level or pseudocode description. Finally, if we want to describe the implementation precisely, we would give the full code of the program (or another fully precise representation, such as in the form of a list of tuples). We call this a formal or low level description.

While we started off by describing NAND-CIRC, NAND-TM, and NAND-RAM programs at the full formal level, as we progress in this book we will move to implementation and high level description. After all, our goal is not to use these models for actual computation, but rather to analyze the general phenomenon of computation. That said, if you don’t understand how the high level description translates to an actual implementation, going “down to the metal” is often an excellent exercise. One of the most important skills for a computer scientist is the ability to move up and down hierarchies of abstractions.

A similar distinction applies to the notion of representation of objects as strings. Sometimes, to be precise, we give a low level specification of exactly how an object maps into a binary string. For example, we
We can describe an algorithm at different levels of granularity/detail and precision. At the highest level we just write the idea in words, omitting all details on representation and implementation. In the intermediate level (also known as implementation or pseudocode) we give enough details of the implementation that would allow someone to derive it, though we still fall short of providing the full code. The lowest level is where the actual code or mathematical description is fully spelled out. These different levels of detail all have their uses, and moving between them is one of the most important skills for a computer scientist.

Figure 7.7: We can describe an algorithm at different levels of granularity/detail and precision. At the highest level we just write the idea in words, omitting all details on representation and implementation. In the intermediate level (also known as implementation or pseudocode) we give enough details of the implementation that would allow someone to derive it, though we still fall short of providing the full code. The lowest level is where the actual code or mathematical description is fully spelled out. These different levels of detail all have their uses, and moving between them is one of the most important skills for a computer scientist.

might describe an encoding of $n$ vertex graphs as length $n^2$ binary strings, by saying that we map a graph $G$ over the vertices $[n]$ to a string $x \in \{0, 1\}^{n^2}$ such that the $n \cdot i + j$-th coordinate of $x$ is 1 if and only if the edge $i \rightarrow j$ is present in $G$. We can also use an intermediate or implementation level description, by simply saying that we represent a graph using the adjacency matrix representation.

Finally, because we are translating between the various representations of graphs (and objects in general) can be done via a NAND-RAM (and hence a NAND-TM) program, when talking in a high level we also suppress discussion of representation altogether. For example, the fact that graph connectivity is a computable function is true regardless of whether we represent graphs as adjacency lists, adjacency matrices, list of edge-pairs, and so on and so forth. Hence, in cases where the precise representation doesn’t make a difference, we would often talk about our algorithms as taking as input an object $X$ (that can be a graph, a vector, a program, etc.) without specifying how $X$ is encoded as a string.

**Defining Algorithms.** Up until now we have use the term “algorithm” informally. However, Turing Machines and the range of equivalent models yield a way to precisely and formally define algorithms. Hence whenever we refer to an algorithm in this book, we will mean that it is an instance of one of the Turing equivalent models, such as Turing machines, NAND-TM, RAM machines, etc. Because of the equivalence of all these models, in many contexts, it will not matter which of these we use.
7.3.3 Turing completeness and equivalence, a formal definition (optional)

A computational model is some way to define what it means for a program (which is represented by a string) to compute a (partial) function. A computational model $M$ is Turing complete, if we can map every Turing machine (or equivalently NAND-TM program) $N$ into a program $P$ for $M$ that computes the same function as $Q$. It is Turing equivalent if the other direction holds as well (i.e., we can map every program in $M$ to a Turing machine that computes the same function). We can define this notion formally as follows. (This formal definition is not crucial for the remainder of this book so feel to skip it as long as you understand the general concept of Turing equivalence; This notion is sometimes referred to in the literature as Gödel numbering or admissable numbering.)

**Definition 7.5 — Turing completeness and equivalence (optional).** Let $\mathcal{F}$ be the set of all partial functions from $\{0, 1\}^*$ to $\{0, 1\}^*$. A computational model is a map $M : \{0, 1\}^* \rightarrow \mathcal{F}$.

We say that a program $P \in \{0, 1\}^*$ $M$-computes a function $F \in \mathcal{F}$ if $M(P) = F$.

A computational model $M$ is Turing complete if there is a computable map $\text{ENCODE}_M : \{0, 1\}^* \rightarrow \{0, 1\}^*$ for every Turing machine $N$ (represented as a string), $M(\text{ENCODE}_M(N))$ is equal to the partial function computed by $P$.

A computational model $M$ is Turing equivalent if it is Turing complete and there exists a computable map $\text{DECODE}_M : \{0, 1\}^* \rightarrow \{0, 1\}^*$ such that or every string $P \in \{0, 1\}^*$, $N = \text{DECODE}_M(P)$ is a string representation of a Turing machine that computes the function $M(P)$.

Some examples of Turing equivalent models (some of which we have already seen, and some are discussed below) include:

- Turing machines
- NAND-TM programs
- NAND-RAM programs
- $\lambda$ calculus
- Game of life (mapping programs and inputs/outputs to starting and ending configurations)
- Programming languages such as Python/C/Javascript/OCaml...
  (allowing for unbounded storage)

### 7.4 CELLULAR AUTOMATA

Many physical systems can be described as consisting of a large number of elementary components that interact with one another. One
way to model such systems is using *cellular automata*. This is a system that consists of a large number (or even infinite) cells. Each cell only has a constant number of possible states. At each time step, a cell updates to a new state by applying some simple rule to the state of itself and its neighbors.

A canonical example of a cellular automaton is *Conway’s Game of Life*. In this automata the cells are arranged in an infinite two dimensional grid. Each cell has only two states: “dead” (which we can encode as 0 and identify with \( \emptyset \)) or “alive” (which we can encode as 1). The next state of a cell depends on its previous state and the states of its 8 vertical, horizontal and diagonal neighbors (see Fig. 7.8). A dead cell becomes alive only if exactly three of its neighbors are alive. A live cell continues to live if it has two or three live neighbors. Even though the number of cells is potentially infinite, we can encode the state using a finite-length string by only keeping track of the live cells. If we initialize the system in a configuration with a finite number of live cells, then the number of live cells will stay finite in all future steps. The Wikipedia page for the Game of Life contains some beautiful figures and animations of configurations that produce very interesting evolutions.

Since the cells in the game of life are are arranged in an infinite two-dimensional grid, it is an example of a *two dimensional cellular automaton*. We can also consider the even simpler setting of a *one dimensional cellular automaton*, where the cells are arranged in an infinite line, see Fig. 7.9. It turns out that even this simple model is enough to achieve Turing-completeness. We will now formally define one-dimensional cellular automata and then prove their Turing completeness.

**Definition 7.6 — One dimensional cellular automata.** Let \( \Sigma \) be a finite set containing the symbol \( \emptyset \). A *one dimensional cellular automaton* over alphabet \( \Sigma \) is described by a transition rule \( r : \Sigma^3 \to \Sigma \), which satisfies \( r(\emptyset, \emptyset, \emptyset) = \emptyset \).
Figure 7.9: In a two dimensional cellular automaton every cell is in position $i, j$ for some integers $i, j \in \mathbb{Z}$. The state of a cell is some value $A_{i,j} \in \Sigma$ for some finite alphabet $\Sigma$. At a given time step, the state of the cell is adjusted according to some function applied to the state of $(i,j)$ and all its neighbors $(i \pm 1, j \pm 1)$. In a one dimensional cellular automaton every cell is in position $i \in \mathbb{Z}$ and the state $A_i$ of $i$ at the next time step depends on its current state and the state of its two neighbors $i - 1$ and $i + 1$.

2 dimensional cellular automaton:

```
\_ | i | i | i | i | i | i | i | / \\
... | 0 | 0 | 0 | 0 | 0 | 0 | 0 | ...
... | 0 | a | c | a | b | a | 0 | ...
... | 0 | b | a | 0 | c | a | 0 | ...
... | 0 | a | b | c | a | b | 0 | ...
... | 0 | 0 | 0 | 0 | 0 | 0 | 0 | ...
```

1 dimensional cellular automaton:

```
... | 0 | a | c | a | b | a | 0 | ...
```

A configuration of the automaton $r$ is a function $A : \mathbb{Z} \to \Sigma$. If an automaton with rule $r$ is in configuration $A$, then its next configuration, denoted by $A' = \text{NEXT}_r(A)$. Is the function $A'$ such that $A'(i) = r(A(i-1), A(i), A(i+1))$ for every $i \in \mathbb{Z}$. In other words, the next state of the automaton $r$ at point $i$ obtained by applying the rule $r$ to the values of $A$ at $i$ and its two neighbors.

Finite configuration. We say that a configuration of an automaton $r$ is finite if there is only some finite number $i_0, \ldots, i_{j-1}$ of indices in $\mathbb{Z}$ such that $A(i_j) \neq \emptyset$. (That is, for every $i \notin \{i_0, \ldots, i_{j-1}\}$, $A(i) = \emptyset$.) Such a configuration can be represented using a finite string that encodes the indices $i_0, \ldots, i_{n-1}$ and the values $A(i_0), \ldots, A(i_{n-1})$. Since $R(\emptyset, \emptyset, \emptyset) = \emptyset$, if $A$ is a finite configuration then $\text{NEXT}_r(A)$ is finite as well. We will only be interested in studying cellular automata that are initialized in finite configurations, and hence remain in a finite configuration throughout their evolution.

7.4.1 One dimensional cellular automata are Turing complete

We can write a program (for example using NAND-RAM) that simulates the evolution of any cellular automaton from an initial finite configuration by simply storing the values of the cells with state not equal to $\emptyset$ and repeatedly applying the rule $r$. Hence cellular automata can be simulated by Turing Machines. What is more surprising that the other direction holds as well. For example, as simple as its rules seem, we can simulate a Turing machine using the game of life (see Fig. 7.10).

In fact, even one dimensional cellular automata can be Turing complete:

Figure 7.10: A Game-of-Life configuration simulating a Turing Machine. Figure by Paul Rendell.
**Theorem 7.7 — One dimensional automata are Turing complete.** For every Turing machine $M$, there is a one dimension cellular automaton that can simulate $M$ on every input $x$.

To make the notion of “simulating a Turing machine” more precise we will need to define configurations of Turing machines. We will do so in Section 7.4.2 below, but at a high level a configuration of a Turing machine is a string that encodes its full state at a given step in its computation. That is, the contents of all (non empty) cells of its tape, its current state, as well as the head position.

The key idea in the proof of Theorem 7.7, that at every point in the computation of a Turing machine $M$, the only cell in $M$’s tape that can change is the one where the head is located, and the value this cell changes to is a function of its current state and the finite state of $M$. This observation allows us to encode the configuration of a Turing machine $M$ as a finite configuration of a cellular automaton $r$, and ensure that a one-step evolution of this encoded configuration under the rules of $r$ corresponds to one step in the execution of the Turing machine $M$.

### 7.4.2 Configurations of Turing machines and the next-step function

To turn the above ideas into a rigorous proof (and even statement!) of Theorem 7.7 we will need precisely define the notion of configurations of Turing machines. This notion will be useful for us in later chapters we well.

**Figure 7.11:** A configuration of a Turing machine $M$ with alphabet $\Sigma$ and state space $[k]$ encodes the state of $M$ at a particular step in its execution as a string $\alpha$ over the alphabet $\Sigma = \Sigma \times (\{\cdot\} \times [k])$. The string is of length $t$ where $t$ is such that $M$’s tape contains $\emptyset$ in all positions $t$ and larger and $M$’s head is in a position smaller than $t$. If $M$’s head is in the $i$-th position, then for $j \neq i$, $\alpha_j$ encodes the value of the $j$-th cell of $M$’s tape, while $\alpha_i$ encodes both this value as well as the current state of $M$. If the machine writes the value $\tau$, changes state to $t$, and moves right, then in the next configuration will contain at position $i$ the value $(\tau, \cdot)$ and at position $i + 1$ the value $(\alpha_{i+1}, t)$.

**Definition 7.8 — Configuration of Turing Machines.** Let $M$ be a Turing machine with tape alphabet $\Sigma$ and state space $[k]$. A configuration of $M$ is a string $\alpha \in \Sigma^*$ where $\Sigma = \Sigma \times (\{\cdot\} \cup [k])$ that satisfies that there
is exactly one coordinate \( i \) for which \( \alpha_i = (\sigma, s) \) for some \( \sigma \in \Sigma \) and \( s \in [k] \). For all other coordinates \( j \), \( \alpha_j = (\sigma', \cdot) \) for some \( \sigma' \in \Sigma \).

A configuration \( \alpha \in \sum' \) of \( M \) corresponds to the following state of its execution:

- \( M \)'s tape contains \( \alpha_{j,0} \) for all \( j < |\alpha| \) and contains \( \emptyset \) for all positions that are at least \( |\alpha| \), where we let \( \alpha_{j,0} \) be the value \( \sigma \) such that \( \alpha_j = (\sigma, t) \) with \( \sigma \in \Sigma \) and \( t \in \{ \cdot \} \cup [k] \). (In other words, since \( \alpha_j \) is a pair of an alphabet symbol \( \sigma \) and either a state in \([k]\) or the symbol \( \cdot \), \( \alpha_{j,0} \) is the first component \( \sigma \) of this pair.)

- \( M \)'s head is in the unique position \( i \) for which \( \alpha_i \) has the form \( (\sigma, s) \) for \( s \in [k] \), and \( M \)'s state is equal to \( s \).

**Definition 7.8** below has some technical details, but is not actually that deep or complicated. Try to take a moment to stop and think how you would encode as a string the state of a Turing machine at a given point in an execution.

Think what are all the components that you need to know in order to be able to continue the execution from this point onwards, and what is a simple way to encode them using a list of finite symbols. In particular, with an eye towards our future applications, try to think of an encoding which will make it as simple as possible to map a configuration at step \( t \) to the configuration at step \( t + 1 \).

**Definition 7.8** is a little cumbersome, but ultimately a configuration is simply a string that encodes a *snapshot* of the state of the NAND-TM program at a given point in the execution. (In operating-systems lingo, it is a “core dump”.) Such a snapshot needs to encode the following components:

1. The current head position.
2. The full contents of the large scale memory, that is the tape.
3. The contents of the “local registers”, that is the state of the machine.

The precise details of how we encode a configuration are not important, but we do want to record the following simple fact:

**Lemma 7.9** Let \( M \) be a Turing machine and let \( NEXT_M : \Sigma^* \rightarrow \Sigma^* \) be the function that maps a configuration of \( M \) to the configuration...
at the next step of the execution. Then for every $i \in \mathbb{N}$, the value of $\text{NEXT}_M(\alpha)_i$ only depends on the coordinates $\alpha_{i-1}, \alpha_i, \alpha_{i+1}$. 2

We leave proving Lemma 7.9 as Exercise 7.6. The idea behind the proof is simple: if the head is neither in position $i$ nor positions $i - 1$ and $i + 1$, then the next-step configuration at $i$ will be the same as it was before. Otherwise, we can “read off” the state of the Turing machine and the value of the tape at the head location from the configuration at $i$ or one of its neighbors and use that to update what the new state at $i$ should be. Completing the full proof is not hard, but doing it is a great way to ensure that you are comfortable with the definition of configurations.

Completing the proof of Theorem 7.7. We can now restate Theorem 7.7 more formally, and complete its proof:

**Theorem 7.10 — One dimensional automata are Turing complete (formal statement).** For every Turing Machine $M$, there is a one-dimensional cellular automaton $r$ over the alphabet $\Sigma^*$ such that

$$E(\text{NEXT}_M(\alpha)) = \text{NEXT}_r(E(\alpha))$$

(7.2)

for every configuration $\alpha \in \Sigma^*$ of $M$, where we denote by $E(\alpha)$ the natural extension of $\alpha$ to a configuration of $r$ obtained by defining $E(\alpha)_i = \alpha_i$ for every $i \in \{0, 1, \ldots, |\alpha| - 1\}$ and $E(\alpha)_j = \emptyset$ for every $j$ such that $j < 0$ or $j \geq |\alpha|$.

**Proof.** We consider the element $(\emptyset, \cdot)$ of $\Sigma$ to correspond to the $\emptyset$ element of the automaton $r$. In this case, by Lemma 7.9, the function $\text{NEXT}_M$ that maps a configuration of $M$ into the next one is in fact a valid rule for a one dimensional automata.

The automaton arising from the proof of Theorem 7.10 has a large alphabet, and furthermore one whose size that depends on the machine $M$ that is being simulated. It turns out that one can obtain an automaton with an alphabet of fixed size that is independent of the program being simulated, and in fact the alphabet of the automaton can be the minimal set $\{0, 1\}$! See Fig. 7.12 for an example of such an Turing-complete automaton.

**Remark 7.11 — Configurations of NAND-TM programs.** We can use the same approach as Definition 7.8 to define configurations of a NAND-TM program. Such a configuration will need to encode:

1. The current value of the variable $i$.
2. For every scalar variable foo, the value of foo.
3. For every array variable Bar, the value Bar[j] for every j ∈ {0, ..., t − 1} where t − 1 is the largest value that the index variable i ever achieved in the computation.

7.5 LAMBDA CALCULUS AND FUNCTIONAL PROGRAMMING LANGUAGES

The λ calculus is another way to define computable functions. It was proposed by Alonzo Church in the 1930’s around the same time as Alan Turing’s proposal of the Turing Machine. Interestingly, while Turing Machines are not used for practical computation, the λ calculus has inspired functional programming languages such as LISP, ML and Haskell, and indirectly the development of many other programming languages as well. In this section we will present the λ calculus and show that its power is equivalent to NAND-TM programs (and hence also to Turing machines). Our Github repository contains a Jupyter notebook with a Python implementation of the λ calculus that you can experiment with to get a better feel for this topic.

The λ operator. At the core of the λ calculus is a way to define “anonymous” functions. For example, instead of giving a name f to a function and defining it as

\[ f(x) = x \times x \]  \hspace{1cm} (7.3)

we can write it as

\[ \lambda x. x \times x \]  \hspace{1cm} (7.4)

and so \((\lambda x. x \times x)(7) = 49\). That is, you can think of \(\lambda x. \text{exp}(x)\), where \text{exp} is some expression as a way of specifying the anonymous function \(x \mapsto \text{exp}(x)\). Anonymous functions, using either \(\lambda x. f(x)\), \(x \mapsto f(x)\) or other closely related notation, appear in many programming languages. For example, in Python we can define the squaring function using \(\lambda \text{a} \text{b} \text{c} : \text{x} \times \text{x} \) while in Javascript we can use \(\lambda \text{x} \mapsto \text{x} \times \text{x}\) or \((\lambda x) \mapsto x \times x\). In Scheme we would define it as \((\lambda \text{a} (\lambda x) (\times \text{a} \text{x}))\). Clearly, the name of the argument to a function doesn’t matter, and so \(\lambda y . y \times y\) is the same as \(\lambda x . x \times x\), as both correspond to the squaring function.

Dropping parenthesis. To reduce notational clutter, when writing λ calculus expressions we often drop the parenthesis for function evaluation. Hence instead of writing \(f(x)\) for the result of applying the function \(f\) to the input \(x\), we can also write this as simply \(f \ x\).
Therefore we can write \((\lambda x. x \times x)7 = 49\). In this chapter, we will use both the \(f(x)\) and \(f x\) notations for function application. Function evaluations are associative and bind from left to right, and hence \(f g h\) is the same as \((f g)h\).

### 7.5.1 Applying functions to functions

A key feature of the \(\lambda\) calculus is that functions are “first-class objects” in the sense that we can use functions as arguments to other functions. For example, can you guess what number is the following expression equal to?

\[
(((\lambda f .(\lambda y.(f (f y))))(\lambda x. x \times x)) 3)
\]

(7.5)

The expression (7.5) might seem daunting, but before you look at the solution below, try to break it apart to its components, and evaluate each component at a time. Working out this example would go a long way toward understanding the \(\lambda\) calculus.

Let’s evaluate (7.5) one step at a time. As nice as it is for the \(\lambda\) calculus to allow anonymous functions, adding names can be very helpful for understanding complicated expressions. So, let us write

\[
F = \lambda f .(\lambda y.(f (f y)))
\]

and

\[
g = \lambda x. x \times x
\]

Therefore (7.5) becomes

\[
((F g) 3).
\]

(7.6)

On input a function \(f\), \(F\) outputs the function \(\lambda y.(f (f y))\), or in other words \(F f\) is the function \(y \mapsto f (f y)\). Our function \(g\) is simply \(g(x) = x^2\) and so \((F g)\) is the function that maps \(y\) to \((y^2)^2 = y^4\). Hence \(((F g)3) = 3^4 = 81\).

**Solved Exercise 7.1** What number does the following expression equal to?

\[
((\lambda x.(\lambda y.x)) 2) 9).
\]

(7.7)

**Solution:**

\(\lambda y.x\) is the function that on input \(y\) ignores its input and outputs \(x\). Hence \((\lambda x.(\lambda y.x))2\) yields the function \(y \mapsto 2\) (or, using \(\lambda\) notation, the function \(\lambda y.2\)). Hence (7.7) is equivalent to \((\lambda y.2) 9 = 2\).
7.5.2 Obtaining multi-argument functions via Currying

In a λ expression of the form λx.e, the expression e can itself involve the λ operator. Thus for example the function

\[ \lambda x. (\lambda y. x + y) \]  \hspace{2cm} (7.8)

maps \( x \) to the function \( y \mapsto x + y \).

In particular, if we invoke the function (7.8) on \( a \) to obtain some function \( f \), and then invoke \( f \) on \( b \), we obtain the value \( a + b \). We can see that the one-argument function (7.8) corresponding to \( a \mapsto (b \mapsto a + b) \) can also be thought of as the two-argument function \( (a, b) \mapsto a + b \). Generally, we can use the λ expression \( \lambda x. (\lambda y. f(x,y)) \) to simulate the effect of a two argument function \( (x, y) \mapsto f(x, y) \). This technique is known as Currying. We will use the shorthand \( \lambda x, y. e \) for \( \lambda x. (\lambda y. e) \). If \( f = \lambda x. (\lambda y. e) \) then \( f(a) b \) corresponds to applying \( f a \) and then invoking the resulting function on \( b \), obtaining the result of replacing in \( e \) the occurrences of \( x \) with \( a \) and occurrences of \( b \) with \( y \). By our rules of associativity, this is the same as \( (f a) b \) which we’ll sometimes also write as \( f(a, b) \).

7.5.3 Formal description of the λ calculus.

We now provide a formal description of the λ calculus. We start with “basic expressions” that contain a single variable such as \( x \) or \( y \) and build more complex expressions of the form \( (e, e') \) and \( \lambda x.e \) where \( e, e' \) are expressions and \( x \) is a variable identifier. Formally λ expressions are defined as follows:

**Definition 7.12 — λ expression.** A λ expression is either a single variable identifier or an expression \( e \) of the one of the following forms:

- **Application:** \( e = (e', e'') \), where \( e' \) and \( e'' \) are λ expressions.
- **Abstraction:** If \( e = \lambda x.(e') \) where \( e' \) is a λ expression.

**Definition 7.12** is a recursive definition since we defined the concept of λ expressions in terms of itself. This might seem confusing at first, but in fact you have known recursive definitions since you were an elementary school student. Consider how we define an arithmetic expression: it is an expression that is either just a number, or has one of the forms \( (e + e'), (e - e'), (e \times e'), \) or \( (e + e') \), where \( e \) and \( e' \) are other arithmetic expressions.

**Free and bound variables.** Variables in a λ expression can either be free or bound to a λ operator (in the sense of Section 1.4.7). In a single-variable λ expression \( \text{var} \), the variable \( \text{var} \) is free. The set of free and bound variables in an application expression \( e = (e', e'') \) is the same as...
as that of the underlying expressions $e'$ and $e''$. In an abstraction expression $e = \lambda \text{var}.(e')$, all free occurrences of var in $e'$ are bound to the $\lambda$ operator of $e$. If you find the notion of free and bound variables confusing, you can avoid all these issues by using unique identifiers for all variables.

**Precedence and parenthesis.** We will use the following rules to allow us to drop some parenthesis. Function application associates from left to right, and so $fgb$ is the same as $(fg)b$. Function application has a higher precedence than the $\lambda$ operator, and so $\lambda x.fgx$ is the same as $\lambda x.(fg)x$. This is similar to how we use the precedence rules in arithmetic operations to allow us to use fewer parenthesis and so write the expression $(7 \times 3) + 2$ as $7 \times 3 + 2$. As mentioned in Section 7.5.2, we also use the shorthand $\lambda x,y.e$ for $\lambda x.(\lambda y.e)$ and the shorthand $f(x,y)$ for $(f \times y)$. This plays nicely with the "Currying" transformation of simulating multi-input functions using $\lambda$ expressions.

**Equivalence of $\lambda$ expressions.** As we have seen in Solved Exercise 7.1, the rule that $(\lambda x.e \exp)f \exp'$ is equivalent to $e \exp'[x \to e \exp']$ enables us to modify $\lambda$ expressions and obtain simpler equivalent form for them. Another rule that we can use is that the parameter does not matter and hence for example $\lambda y.y$ is the same as $\lambda z.z$. Together these rules define the notion of equivalence of $\lambda$ expressions:

**Definition 7.13 — Equivalence of $\lambda$ expressions.** Two $\lambda$ expressions are **equivalent** if they can be made into the same expression by repeated applications of the following rules:

1. **Evaluation (aka $\beta$ reduction):** The expression $(\lambda x.e \exp)f \exp'$ is equivalent to $e \exp'[x \to f \exp']$.

2. **Variable renaming (aka $\alpha$ conversion):** The expression $\lambda x.e \exp$ is equivalent to $\lambda y.e \exp[x \to y]$.

If $e \exp$ is a $\lambda$ expression of the form $\lambda x.e \exp'$ then it naturally corresponds to the function that maps any input $z$ to $e \exp'[x \to z]$. Hence the $\lambda$ calculus naturally implies a computational model. Since in the $\lambda$ calculus the inputs can themselves be functions, we need to decide in what order we evaluate an expression such as

$$(\lambda x.f)(\lambda y.gz) . \quad (7.9)$$

There are two natural conventions for this:

- **Call by name** (aka "lazy evaluation"): We evaluate $(7.9)$ by first plugging in the righthand expression $(\lambda y.gz)$ as input to the lefthand
side function, obtaining \( f[x \to (\lambda y.g z)] \) and then continue from there.

- Call by value (aka “eager evaluation”): We evaluate (7.9) by first evaluating the righthand side and obtaining \( h = g[y \to z] \), and then plugging this into the lefthandside to obtain \( f[x \to h] \).

Because the \( \lambda \) calculus has only pure functions, that do not have “side effects”, in many cases the order does not matter. In fact, it can be shown that if we obtain an definite irreducible expression (for example, a number) in both strategies, then it will be the same one. However, for concreteness we will always use the “call by name” (i.e., lazy evaluation) order. (The same choice is made in the programming language Haskell, though many other programming languages use eager evaluation.) Formally, the evaluation of a \( \lambda \) expression using “call by name” is captured by the following process:

**Definition 7.14 — Simplification of \( \lambda \) expressions.** Let \( e \) be a \( \lambda \) expression. The simplification of \( e \) is the result of the following recursive process:

1. If \( e \) is a single variable \( x \) then the simplification of \( e \) is \( e \).

2. If \( e \) has the form \( e = \lambda x.e' \) then the simplification of \( e \) is \( \lambda x.f' \) where \( f' \) is the simplification of \( f' \).

3. (Evaluation / \( \beta \) reduction.) If \( e \) has the form \( e = (\lambda x.e'e'') \) then the simplification of \( e \) is the simplification of \( e'[x \to e''] \), which denotes replacing all copies of \( x \) in \( e' \) bound to the \( \lambda \) operator with \( e'' \).

4. (Renaming / \( \alpha \) conversion.) The canonical simplification of \( e \) is obtained by taking the simplification of \( e \) and renaming the variables so that the first bound variable in the expression is \( v_0 \), the second one is \( v_1 \), and so on and so forth.

We say that two \( \lambda \) expressions \( e \) and \( e' \) are equivalent, denoted by \( e \equiv e' \), if they have the same canonical simplification.

**Solved Exercise 7.2 — Equivalence of \( \lambda \) expressions.** Prove that the following two expressions \( e \) and \( f \) are equivalent:

\[
e = \lambda x.x
\]

(7.10)

\[
f = (\lambda a.(\lambda b.b))(\lambda z.zz)
\]

(7.11)
7.5.4 Infinite loops in the $\lambda$ calculus

Like Turing machines and NAND-TM programs, the simplification process in the $\lambda$ calculus can also enter into an infinite loop. For example, consider the $\lambda$ expression

$$\lambda x.xx \lambda x.xx$$

If we try to simplify (7.12) by invoking the lefthand function on the righthand one, then we get another copy of (7.12) and hence this never ends. There are examples where the order of evaluation can matter for whether or not an expression can be simplified, see Exercise 7.8.

7.6 THE “ENHANCED” $\lambda$ CALCULUS

We now discuss the $\lambda$ calculus as a computational model. We will start by describing an “enhanced” version of the $\lambda$ calculus that contains some “superfluous features” but is easier to wrap your head around. We will first show how the enhanced $\lambda$ calculus is equivalent to Turing machines in computational power. Then we will show how all the features of “enhanced $\lambda$ calculus” can be implemented as “syntactic sugar” on top of the “pure” (i.e., non enhanced) $\lambda$ calculus. Hence the pure $\lambda$ calculus is equivalent in power to Turing machines (and hence also to RAM machines and all other Turing-equivalent models).

The enhanced $\lambda$ calculus includes the following set of objects and operations:

- **Boolean constants and IF function**: There are $\lambda$ expressions $0$, $1$ and $\text{IF}$ that satisfy the following conditions: for every $\lambda$ expression $e$ and $f$, $\text{IF} \ 1 \ e \ f = e$ and $\text{IF} \ 0 \ e \ f = f$. That is, $\text{IF}$ is the function that given three arguments $a$, $e$, $f$ outputs $e$ if $a = 1$ and $f$ if $a = 0$.

- **Pairs**: There is a $\lambda$ expression $\text{PAIR}$ which we will think of as the pairing function. For every $\lambda$ expressions $e$, $f$ $\text{PAIR} \ e \ f$ if the pair $(e, f)$ that contains $e$ as its first member and $f$ as its second member. We also have $\lambda$ expressions $\text{HEAD}$ and $\text{TAIL}$ that extract the first and second member of a pair respectively. Hence, for every $\lambda$ expressions $e$, $f$, $\text{HEAD} (\text{PAIR} \ e \ f) = e$ and $\text{TAIL} (\text{PAIR} \ e \ f) = f$.

3 In Lisp, the $\text{PAIR}$, $\text{HEAD}$ and $\text{TAIL}$ functions are traditionally called cons, car and cdr.
• **Lists and strings:** There is λ expression $\text{NIL}$ that corresponds to the empty list, which we also denote by $\langle \bot \rangle$. Using $\text{PAIR}$ and $\text{NIL}$ we construct lists. The idea is that if $L$ is a $k$ element list of the form $\langle e_1, e_2, \ldots, e_k, \bot \rangle$ then for every λ expression $e_0$ we can obtain the $k + 1$ element list $\langle e_0, e_1, e_2, \ldots, e_k, \bot \rangle$ using the expression $\text{PAIR} e_0 L$.

For example, for every three λ expressions $e, f, g$, the following corresponds to the three element list $\langle e, f, g, \bot \rangle$:

$$\text{PAIR} e (\text{PAIR} f (\text{PAIR} g \text{NIL})).$$  \hspace{1cm} (7.13)

The λ expression $\text{ISEMPTY}$ returns 1 on $\text{NIL}$ and returns 0 on every other list. A string is simply a list of bits.

• **List operations:** The enhanced λ calculus also contains the list-processing functions $\text{MAP}$, $\text{REDUCE}$, and $\text{FILTER}$. Given a list $L = \langle x_0, \ldots, x_{n-1}, \bot \rangle$ and a function $f$, $\text{MAP} L f$ applies $f$ on every member of the list to obtain the new list $L' = \langle f(x_0), \ldots, f(x_{n-1}), \bot \rangle$. Given a list $L$ as above and an expression $f$ whose output is either 0 or 1, $\text{FILTER} L f$ returns the list $\langle x_i \rangle_{f(x_i) = 1}$ containing all the elements of $L$ for which $f$ outputs 1. The function $\text{REDUCE}$ applies a “combining” operation to a list. For example, $\text{REDUCE} L + 0$ will return the sum of all the elements in the list $L$. More generally, $\text{REDUCE}$ takes a list $L$, an operation $f$ (which we think of as taking two arguments) and a λ expression $z$ (which we think of as the “neutral element” for the operation $f$, such as 0 for addition and 1 for multiplication). The output is defined via

$$\text{REDUCE} L f z = \begin{cases} z & L = \text{NIL} \\ f (\text{HEAD} L) (\text{REDUCE} (\text{TAIL} L) f z) & \text{otherwise} \end{cases}.$$ \hspace{1cm} (7.14)

See Fig. 7.15 for an illustration of the three list-processing operations.

• **Recursion:** Finally, we want to be able to execute recursive functions. Since in λ calculus functions are anonymous, we can’t write a definition of the form $f(x) = \text{blah}$ where $\text{blah}$ includes calls to $f$. Instead we use functions $f$ that take an additional input $\text{me}$ as a parameter. The operator $\text{RECURSE}$ will take such a function $f$ as input and return a “recursive version” of $f$ where all the calls to $\text{me}$ are replaced by recursive calls to this function. That is, if we have a function $F$ taking two parameters $\text{me}$ and $x$, then $\text{RECURSE} F$ will be the function $f$ taking one parameter $x$ such that $f(x) = F(f, x)$ for every $x$. 
Solved Exercise 7.3 — Compute NAND using λ calculus. Give a λ expression $N$ such that $N \ x \ y = \text{NAND}(x, y)$ for every $x, y \in \{0, 1\}$.

Solution:

The NAND of $x, y$ is equal to 1 unless $x = y = 1$. Hence we can write

$$N = \lambda x, y. \text{IF}(x, \text{IF}(y, 0, 1), 1) \quad (7.15)$$

Solved Exercise 7.4 — Compute XOR using λ calculus. Give a λ expression XOR such that for every list $L = \langle x_0, ..., x_{n-1}, \bot \rangle$ where $x_i \in \{0, 1\}$ for $i \in [n]$, XOR$L$ evaluates to $\sum x_i \mod 2$.

Solution:

First, we note that we can compute XOR of two bits as follows:

$$\text{NOT} = \lambda a. \text{IF}(a, 0, 1) \quad (7.16)$$

and

$$\text{XOR}_2 = \lambda a, b. \text{IF}(b, \text{NOT}(a), a) \quad (7.17)$$

(We are using here a bit of syntactic sugar to describe the functions. To obtain the λ expression for XOR we will simply replace the expression (7.16) in (7.17).) Now recursively we can define the XOR of a list as follows:

$$\text{XOR}(L) = \begin{cases} 0 & \text{L is empty} \\ \text{XOR}_2(\text{HEAD}(L), \text{XOR}(\text{TAIL}(L))) & \text{otherwise} \end{cases} \quad (7.18)$$

This means that XOR is equal to

$$\text{RECURSE} \ (\lambda me, L. \text{IF}(\text{ISEMPTY}(L), 0, \text{XOR}_2(\text{HEAD}(L), me(\text{TAIL}(L))))). \quad (7.19)$$

That is, XOR is obtained by applying the RECURSE operator to the function that on inputs $me, L$, returns 0 if ISEMPTY$(L)$ and otherwise returns XOR$_2$ applied to HEAD$(L)$ and to $me(TAIL(L))$.

We could have also computed XOR using the REDUCE operation, we leave working this out as an exercise to the reader.
7.6.1 Computing a function in the enhanced λ calculus

An enhanced λ expression is obtained by composing the objects above with the application and abstraction rules. The result of simplifying a λ expression is an equivalent expression, and hence if two expressions have the same simplification then they are equivalent.

**Definition 7.15 — Computing a function via λ calculus.** Let $F : \{0, 1\}^* \rightarrow \{0, 1\}^*$

We say that $\exp$ computes $F$ if for every $x \in \{0, 1\}^*$,

$$\exp(x_0, \ldots, x_{n-1}, \perp \langle \cong \langle y_0, \ldots, y_{m-1}, \perp \rangle \rangle (7.20)$$

where $n = |x|$, $y = F(x)$, and $m = |y|$, and the notion of equivalence is defined as per Definition 7.14.

7.6.2 Enhanced λ calculus is Turing-complete

The basic operations of the enhanced λ calculus more or less amount to the Lisp or Scheme programming languages. Given that, it is perhaps not surprising that the enhanced λ-calculus is equivalent to Turing machines:

**Figure 7.14:** A list $(x_0, x_1, x_2)$ in the λ calculus is constructed from the tail up, building the pair $(x_2, \NIL)$, then the pair $(x_1, (x_2, \NIL))$ and finally the pair $(x_0, (x_1, (x_2, \NIL)))$. That is, a list is a pair where the first element of the pair is the first element of the list and the second element is the rest of the list. The figure on the left renders this “pairs inside pairs” construction, though it is often easier to think of a list as a “chain”, as in the figure on the right, where the second element of each pair is thought of as a link, pointer or reference to the remainder of the list.

**Figure 7.15:** Illustration of the MAP, FILTER and REDUCE operations.
Theorem 7.16 — Lambda calculus and NAND-TM. For every function $F : \{0, 1\}^* \to \{0, 1\}^*$, $F$ is computable in the enhanced $\lambda$ calculus if and only if it is computable by a Turing machine.

Proof Idea:

To prove the theorem, we need to show that (1) if $F$ is computable by a $\lambda$ calculus expression then it is computable by a Turing machine, and (2) if $F$ is computable by a Turing machine, then it is computable by an enhanced $\lambda$ calculus expression.

Showing (1) is fairly straightforward. Applying the simplification rules to a $\lambda$ expression basically amounts to “search and replace” which we can implement easily in, say, NAND-RAM, or for that matter Python (both of which are equivalent to Turing machines in power). Showing (2) essentially amounts to simulating a Turing machine (or writing a NAND-TM interpreter) in a functional programming language such as LISP or Scheme. We give the details below but how this can be done is a good exercise in mastering some functional programming techniques that are useful in their own right.

Proof of Exercise 7.10. We only sketch the proof. The “if” direction is simple. As mentioned above, evaluating $\lambda$ expressions basically amounts to “search and replace”. It is also a fairly straightforward programming exercise to implement all the above basic operations in an imperative language such as Python or C, and using the same ideas we can do so in NAND-RAM as well, which we can then transform to a NAND-TM program.

For the “only if” direction we need to simulate a Turing machine using a $\lambda$ expression. We will do so by first showing that showing for every Turing machine $M$ a $\lambda$ expression to compute the next-step function $\text{NEXT}_M : \Sigma^* \to \Sigma^*$ that maps a configuration of $M$ to the next one (see Section 7.4.2).

A configuration of $M$ is a string $\alpha \in \Sigma^*$ for a finite set $\Sigma$. We can encode every symbol $\sigma \in \Sigma$ by a finite string $\{0, 1\}^\ell$, and so we will encode a configuration $\alpha$ in the $\lambda$ calculus as a list $\langle \alpha_0, \alpha_1, \ldots, \alpha_{m-1}, \perp \rangle$ where $\alpha_i$ is an $\ell$-length string (i.e., an $\ell$-length list of 0’s and 1’s) encoding a symbol in $\Sigma$.

By Lemma 7.9, for every $\alpha \in \Sigma^*$, $\text{NEXT}_M(\alpha)_i$ is equal to $r(\alpha_{i-1}, \alpha_i, \alpha_{i+1})$ for some finite function $r : \Sigma^3 \to \Sigma$. Using our encoding of $\Sigma$ as $\{0, 1\}^\ell$, we can also think of $r$ as mapping $\{0, 1\}^{3\ell}$ to $\{0, 1\}^\ell$. By Solved Exercise 7.3, we can compute the NAND function, and hence every finite function, including $r$, using the $\lambda$ calculus. Using this insight, we can compute $\text{NEXT}_M$ using the $\lambda$ calculus as
follows. Given a list \( L \) encoding the configuration \( \alpha_0 \cdots \alpha_{m−1} \), we define the lists \( L_{\text{prev}} \) and \( L_{\text{next}} \) encoding the configuration \( \alpha \) shifted by one step to the right and left respectively. The next configuration \( \alpha' \) is defined as \( \alpha'_i = r(L_{\text{prev}}[i], L[i], L_{\text{next}}[i]) \) where we let \( L'[i] \) denote the \( i \)-th element of \( L' \). This can be computed by recursion (and hence using the enhanced \( \lambda \) calculus' \texttt{RECURSE} operator) as follows:

Algorithm 7.17 — \( NEXTh \) using the \( \lambda \) calculus.

Input: List \( L = \langle \alpha_0, \alpha_1, \ldots, \alpha_{m−1}, \bot \rangle \) encoding a configuration \( \alpha \).

Output: List \( L' \) encoding \( NEXTh(\alpha) \)

1: procedure \texttt{ComputeNext}(\( L_{\text{prev}} \), \( L \), \( L_{\text{next}} \))
2: if \( IS\text{EMPTY} \) \( L_{\text{prev}} \)
3: return \texttt{NIL}
4: end if
5: \( a \leftarrow \text{HEAD} L_{\text{prev}} \)
6: if \( IS\text{EMPTY} \) \( L \)
7: \( b \leftarrow \emptyset \) \# Encoding of \( \emptyset \) in \{0, 1\}^L
8: else
9: \( b \leftarrow \text{HEAD} L \)
10: end if
11: if \( IS\text{EMPTY} \) \( L_{\text{next}} \)
12: \( c \leftarrow \emptyset \)
13: else
14: \( c \leftarrow \text{HEAD} L_{\text{next}} \)
15: end if
16: return \( \text{PAIR} \ r(a, b, c) \ \texttt{ComputeNext}(\text{TAIL} \ L_{\text{prev}}, \ \text{TAIL} \ L, \ \text{TAIL} \ L_{\text{next}}) \)
17: end procedure
18: \( L_{\text{prev}} \leftarrow \text{PAIR} \emptyset \ L \) \# \( L_{\text{prev}} = \langle \emptyset, \alpha_0, \ldots, \alpha_{m−1}, \bot \rangle \)
19: \( L_{\text{next}} \leftarrow \text{TAIL} L \) \# \( L_{\text{next}} = \langle \alpha_1, \ldots, \alpha_{m−1}, \bot \rangle \)
20: return \texttt{ComputeNext}(\( L_{\text{prev}} \), \( L \), \( L_{\text{next}} \))

Once we can compute \( NEXTh \), we can simulate the execution of \( M \) on input \( x \) using the following recursion. Define \( \text{FINAL}(\alpha) \) to be the final configuration of \( M \) when initialized at configuration \( \alpha \). The function \( \text{FINAL} \) can be defined recursively as follows:

\[
\text{FINAL}(\alpha) = \begin{cases} 
\alpha & \text{\( \alpha \) is halting configuration} \\
\text{\( NEXTh(\alpha) \)} & \text{otherwise}
\end{cases}
\] (7.21)

Checking whether a configuration is halting (i.e., whether it is one in which the transition function would output \( \text{Halt} \)) can be easily implemented in the \( \lambda \) calculus, and hence we can use the \texttt{RECURSE} to compute \( \text{FINAL} \). If we let \( \alpha^0 \) be the initial configuration of \( M \) on
input \( x \) then we can obtain the output \( M(x) \) from \( \text{FINAL}(\alpha^0) \), hence completing the proof.

\[
\]

### 7.7 FROM ENHANCED TO PURE \( \lambda \) CALCULUS

While the collection of “basic” functions we allowed for the enhanced \( \lambda \) calculus is smaller than what’s provided by most Lisp dialects, coming from NAND-TM it still seems a little “bloated”. Can we make do with less? In other words, can we find a subset of these basic operations that can implement the rest?

It turns out that there is in fact a proper subset of the operations of the enhanced \( \lambda \) calculus that can be used to implement the rest. That subset is the empty set. That is, we can implement all the operations above using the \( \lambda \) formalism only, even without using 0’s and 1’s. It’s \( \lambda \)'s all the way down!

---

This is a good point to pause and think how you would implement these operations yourself. For example, start by thinking how you could implement MAP using REDUCE, and then REDUCE using RECURSE combined with 0, 1, IF, PAIR, HEAD, TAIL, NIL, ISEMPTY. You can also PAIR, HEAD and TAIL based on 0, 1, IF. The most challenging part is to implement RECURSE using only the operations of the pure \( \lambda \) calculus.

---

**Theorem 7.18 — Enhanced \( \lambda \) calculus equivalent to pure \( \lambda \) calculus.** There are \( \lambda \) expressions that implement the functions \( 0, 1, \text{IF}, \text{PAIR}, \text{HEAD}, \text{TAIL}, \text{NIL}, \text{ISEMPTY}, \text{MAP}, \text{REDUCE}, \) and RECURSE.

The idea behind **Theorem 7.18** is that we encode 0 and 1 themselves as \( \lambda \) expressions, and build things up from there. This is known as **Church encoding**, as it was originated by Church in his effort to show that the \( \lambda \) calculus can be a basis for all computation. We will not write the full formal proof of **Theorem 7.18** but outline the ideas involved in it:

- We define 0 to be the function that on two inputs \( x, y \) outputs \( y \), and 1 to be the function that on two inputs \( x, y \) outputs \( x \). We use Currying to achieve the effect of two-input functions and hence \( 0 = \lambda x.\lambda y. y \) and \( 1 = \lambda x.\lambda y. x \). (This representation scheme is the common convention for representing false and true but there are many other alternative representations for 0 and 1 that would have worked just as well.)
• The above implementation makes the \textit{IF} function trivial: 
\[\text{IF}(\text{cond}, a, b) \] is simply \textit{cond} \textit{a} \textit{b} since \textit{0a} = \textit{b} and \textit{1a} = \textit{a}. We can write \textit{IF} = \lambda x. x to achieve \textit{IF}(\text{cond}, a, b) = ((\text{IF} \text{cond}) a) b = \text{cond} a b.

• To encode a pair \((x, y)\) we will produce a function \(f_{x,y}\) that has \(x\) and \(y\) “in its belly” and satisfies \(f_{x,y} g = gx y\) for every function \(g\). That is, \(\text{PAIR} = \lambda x, y. (\lambda g. g x y)\). We can extract the first element of a pair \(p\) by writing \(p1\) and the second element by writing \(p0\), and so \(\text{HEAD} = \lambda p. p1\) and \(\text{TAIL} = \lambda p. p0\).

• We define \textit{NIL} to be the function that ignores its input and always outputs 1. That is, \textit{NIL} = \lambda x. 1. The \textit{ISEMPTY} function checks, given an input \(p\), whether we get 1 if we apply \(p\) to the function \(\text{zero} = \lambda x, y. 0\) that ignores both its inputs and always outputs 0. For every valid pair of the form \(p = \text{PAIR} xy\), \(p\text{zero} = p xy = 0\) while \(\text{NIL} zero = 1\). Formally, \(\text{ISEMPTY} = \lambda p. p(\lambda x, y. 0)\).

\begin{remark}
\textbf{Remark 7.19 — Church numerals (optional).} There is nothing special about Boolean values. You can use similar tricks to implement natural numbers using \(\lambda\) terms. The standard way to do so is to represent the number \(n\) by the function \(\text{ITER}_n\) that on input a function \(f\) outputs the function \(x \mapsto f(f(\ldots f(x)))\) \((n\ \text{times})\). That is, we represent the natural number 1 as \(\lambda f. f\), the number 2 as \(\lambda f. (\lambda x. f(fx))\), the number 3 as \(\lambda f. (\lambda x. f(f(fx)))\), and so on and so forth. (Note that this is not the same representation we used for 1 in the Boolean context: this is fine; we already know that the same object can be represented in more than one way.) The number 0 is represented by the function that maps any function \(f\) to the identity function \(\lambda x. x\). (That is, \(0 = \lambda f. (\lambda x. x)\).)
In this representation, we can compute \(\text{PLUS}(n, m)\) as \(\lambda f. \lambda x. (n f)((m f)x)\) and \(\text{TIMES}(n, m)\) as \(\lambda f. n(m f)\). Subtraction and division are trickier, but can be achieved using recursion. (Working this out is a great exercise.)
\end{remark}

7.7.1 List processing
Now we come to a bigger hurdle, which is how to implement \textit{MAP}, \textit{FILTER}, \textit{REDUCE} and \textit{RECURSE} in the pure \(\lambda\) calculus. It turns out that we can build \textit{MAP} and \textit{FILTER} from \textit{REDUCE}, and \textit{REDUCE} from \textit{RECURSE}. For example \(\text{MAP}(L, f)\) is the same as \(\text{REDUCE}(L, g)\) where \(g\) is the operation that on input \(x\) and \(y\), outputs \(\text{PAIR}(f(x), \text{NIL})\) if \(y\) is \text{NIL} and otherwise outputs \(\text{PAIR}(f(x), y)\). (I leave checking this as a (recommended!) exercise for you, the reader.)
We can define REDUCE($L$, $g$) recursively, by setting
REDUCE($\text{NIL}$, $g$) = $\text{NIL}$ and stipulating that given a non-empty list $L$, which we can think of as a pair ($\text{head}$, $\text{rest}$),
REDUCE($L$, $g$) = $g(\text{head}, \text{REDUCE}(\text{rest}, g))$. Thus, we might try to write a recursive $\lambda$ expression for REDUCE as follows

\[
\text{REDUCE} = \lambda L, g. \text{IF}(\text{ISEMPTY}(L), \text{NIL}, g\text{HEAD}(L)\text{REDUCE}(\text{TAIL}(L), g)) .
\]

(7.22)

The only fly in this ointment is that the $\lambda$ calculus does not have the notion of recursion, and so this is an invalid definition. But of course we can use our RECURSE operator to solve this problem. We will replace the recursive call to “REDUCE” with a call to a function $me$ that is given as an extra argument, and then apply RECURSE to this. Thus $\text{REDUCE} = \text{RECURSE myREDUCE}$ where

\[
\text{myREDUCE} = \lambda me, L, g. \text{IF}(\text{ISEMPTY}(L), \text{NIL}, g\text{HEAD}(L)me(\text{TAIL}(L), g)) .
\]

(7.23)

7.7.2 The $Y$ combinator, or recursion without recursion

Eq. (7.23) means that implementing MAP, FILTER, and REDUCE boils down to implementing the RECURSE operator in the pure $\lambda$ calculus. This is what we do now.

How can we implement recursion without recursion? We will illustrate this using a simple example - the XOR function. As shown in Solved Exercise 7.4, we can write the XOR function of a list recursively as follows:

\[
\text{XOR}(L) = \begin{cases} 
0 & \text{L is empty} \\
\text{XOR}_2(\text{HEAD}(L), \text{XOR}(\text{TAIL}(L))) & \text{otherwise}
\end{cases}
\]

(7.24)

where $\text{XOR}_2 : \{0, 1\}^2 \rightarrow \{0, 1\}$ is the XOR on two bits. In Python we would write this as

```python
def xor2(a,b): return 1-b if a else b
def head(L): return L[0]
def tail(L): return L[1:]

def xor(L): return xor2(head(L),xor(tail(L))) if L else 0

print(xor([0,1,1,0,0,1]))
# 1
```

Now, how could we eliminate this recursive call? The main idea is that since functions can take other functions as input, it is perfectly
legal in Python (and the λ calculus of course) to give a function itself as input. So, our idea is to try to come up with a non recursive function tempxor that takes two inputs: a function and a list, and such that tempxor(tempxor,L) will output the XOR of L!

At this point you might want to stop and try to implement this on your own in Python or any other programming language of your choice (as long as it allows functions as inputs).

Our first attempt might be to simply use the idea of replacing the recursive call by me. Let’s define this function as myxor

```python
def myxor(me,L):
    return xor2(head(L),me(tail(L))) if L else 0
```

Let’s test this out:

```python
myxor(myxor,[1,0,1])
```

If you do this, you will get the following complaint from the interpreter:

```
TypeError: myxor() missing 1 required positional argument
```

The problem is that myxor expects two inputs- a function and a list- while in the call to me we only provided a list. To correct this, we modify the call to also provide the function itself:

```python
def tempxor(me,L):
    return xor2(head(L),me(me,tail(L))) if L else 0
```

Note the call me(me,..) in the definition of tempxor: given a function me as input, tempxor will actually call the function me with itself as the first input. If we test this out now, we see that we actually get the right result!

```python
tempxor(tempxor,[1,0,1])
# 0
tempxor(tempxor,[1,0,1,1])
# 1
```

and so we can define xor(L) as simply return tempxor(tempxor,L).

The approach above is not specific to XOR. Given a recursive function f that takes an input x, we can obtain a non recursive version as follows:
1. Create the function \texttt{myf} that takes a pair of inputs \texttt{me} and \texttt{x}, and replaces recursive calls to \texttt{f} with calls to \texttt{me}.

2. Create the function \texttt{tempf} that converts calls in \texttt{myf} of the form \texttt{me(x)} to calls of the form \texttt{me(me,x)}.

3. The function \texttt{f(x)} will be defined as \texttt{tempf(tempf,x)}

Here is the way we implement the \texttt{RECURSE} operator in Python. It will take a function \texttt{myf} as above, and replace it with a function \texttt{g} such that \texttt{g(x)=myf(g,x)} for every \texttt{x}.

\begin{verbatim}
def RECURSE(myf):
    def tempf(me,x):
        return myf(lambda y: me(me,y),x)
    return lambda x: tempf(tempf,x)

xor = RECURSE(myxor)
print(xor([0,1,0,0,1]))
# 1
print(xor([1,1,0,0,1,1,1,1]))
# 0
\end{verbatim}

\textbf{From Python to the calculus.} In the \(\lambda\) calculus, a two input function \(g\) that takes a pair of inputs \texttt{me, y} is written as \(\lambda me.(\lambda y.g)\). So the function \(y \mapsto me(me, y)\) is simply written as \texttt{me me} and similarly the function \(x \mapsto tempf(tempf, x)\) is simply \texttt{tempf tempf}. (Can you see why?) Therefore the function \texttt{tempf} defined above can be written as \(\lambda me. myf(me me)\). This means that if we denote the input of \texttt{RECURSE} by \texttt{f}, then \texttt{RECURSE myf = tempf tempf} where \texttt{tempf = \lambda m.f(m m)} or in other words

\[
\text{RECURSE} = \lambda f.((\lambda m.f(m m)) (\lambda m.f(m m))) \quad (7.25)
\]

The \texttt{online appendix} contains an implementation of the \(\lambda\) calculus using Python. Here is an implementation of the recursive XOR function from that appendix:\footnote{Because of specific issues of Python syntax, in this implementation we use \texttt{f * g} for applying \texttt{f} to \texttt{g} rather than \texttt{fg}, and use \texttt{\lambda x.exp} rather than \texttt{\lambda x.exp} for abstraction. We also use \texttt{\_0} and \texttt{\_1} for the \(\lambda\) terms for \texttt{0} and \texttt{1} so as not to confuse with the Python constants.}

\begin{verbatim}
# XOR of two bits
XOR2 = \lambda(a,b)(IF(a,IF(b,\_0,\_1),b))

# Recursive XOR with recursive calls replaced by m
\_m parameter
myXOR = \lambda(m,l)(IF(ISEMPTY(l),\_0,XOR2(HEAD(l),m(TAIL(l)))))
\end{verbatim}
# Recurse operator (aka Y combinator)
RECURSE = \lambda f((\lambda m(f(m*m)))(\lambda m(f(m*m))))

# XOR function
XOR = RECURSE(myXOR)

#TESTING:
XOR(PAIR(_1,NIL)) # List [1]  
# equals 1

XOR(PAIR(_1,PAIR(_0,PAIR(_1,NIL)))) # List [1,0,1]  
# equals 0

**Remark 7.20 — The Y combinator.** The RECURSE operator above is better known as the Y combinator. It is one of a family of fixed point operators that given a lambda expression \( F \), find a fixed point \( f \) of \( F \) such that \( f = Ff \). If you think about it, XOR is the fixed point of \( myXOR \) above. XOR is the function such that for every \( x \), if plug in XOR as the first argument of \( myXOR \) then we get back XOR, or in other words \( XOR = myXOR \cdot XOR \). Hence finding a fixed point for \( myXOR \) is the same as applying RECURSE to it.

7.8 THE CHURCH-TURING THESIS (DISCUSSION)

“[In 1934], Church had been speculating, and finally definitely proposed, that the \( \lambda \)-definable functions are all the effectively calculable functions .... When Church proposed this thesis, I sat down to disprove it ... but, quickly realizing that [my approach failed], I became overnight a supporter of the thesis.”, Stephen Kleene, 1979.

“[The thesis is] not so much a definition or an axiom but ... a natural law.”, Emil Post, 1936.

We have defined functions to be computable if they can be computed by a NAND-TM program, and we’ve seen that the definition would remain the same if we replaced NAND-TM programs by Python programs, Turing machines, \( \lambda \) calculus, cellular automata, and many other computational models. The Church-Turing thesis is that this is the only sensible definition of “computable” functions. Unlike the
“Physical Extended Church Turing Thesis” (PECTT) which we saw before, the Church Turing thesis does not make a concrete physical prediction that can be experimentally tested, but it certainly motivates predictions such as the PECTT. One can think of the Church-Turing Thesis as either advocating a definitional choice, making some prediction about all potential computing devices, or suggesting some laws of nature that constrain the natural world. In Scott Aaronson’s words, “whatever it is, the Church-Turing thesis can only be regarded as extremely successful”. No candidate computing device (including quantum computers, and also much less reasonable models such as the hypothetical “closed time curve” computers we mentioned before) has so far mounted a serious challenge to the Church Turing thesis. These devices might potentially make some computations more efficient, but they do not change the difference between what is finitely computable and what is not. (The extended Church Turing thesis, which we discuss in Section 12.3, stipulates that Turing machines capture also the limit of what can be efficiently computable. Just like its physical version, quantum computing presents the main challenge to this thesis.)

7.8.1 Different models of computation
We can summarize the models we have seen in the following table:

<table>
<thead>
<tr>
<th>Computational problems</th>
<th>Type of model</th>
<th>Examples</th>
</tr>
</thead>
<tbody>
<tr>
<td>Finite functions $f : {0,1}^n \to {0,1}^m$</td>
<td>Non uniform computation (algorithm depends on input length)</td>
<td>Boolean circuits, NAND circuits, straight-line programs (e.g., NAND-CIRC)</td>
</tr>
<tr>
<td>Functions with unbounded inputs $F : {0,1}^* \to {0,1}^*$</td>
<td>Sequential access to memory</td>
<td>Turing machines, NAND-TM programs</td>
</tr>
<tr>
<td>–</td>
<td>Indexed access / RAM</td>
<td>RAM machines, NAND-RAM, modern programming languages</td>
</tr>
<tr>
<td>–</td>
<td>Other</td>
<td>Lambda calculus, cellular automata</td>
</tr>
</tbody>
</table>
Later on in Chapter 16 we will study memory bounded computation. It turns out that NAND-TM programs with a constant amount of memory are equivalent to the model of finite automata (the adjectives “deterministic” or “nondeterministic” are sometimes added as well, this model is also known as finite state machines) which in turns captures the notion of regular languages (those that can be described by regular expressions), which is a concept we will see in Chapter 9.

### Lecture Recap

- While we defined computable functions using Turing machines, we could just as well have done so using many other models, including not just NAND-TM programs but also RAM machines, NAND-RAM, the λ-calculus, cellular automata and many other models.
- Very simple models turn out to be “Turing complete” in the sense that they can simulate arbitrarily complex computation.

### 7.9 Exercises

**Exercise 7.1 — NAND-TM lookup.** This exercise shows part of the proof that NAND-TM can simulate NAND-RAM. Produce the code of a NAND-TM program that computes the function $\text{LOOKUP} : \{0, 1\}^* \rightarrow \{0, 1\}$ that is defined as follows. On input $pf(i)x$, where $pf(i)$ denotes a prefix-free encoding of an integer $i$, $\text{LOOKUP}(pf(i)x) = x_i$ if $i < |x|$ and $\text{LOOKUP}(pf(i)x) = 0$ otherwise. (We don’t care what $\text{LOOKUP}$ outputs on inputs that are not of this form.) You can choose any prefix-free encoding of your choice, and also can use your favorite programming language to produce this code.

**Exercise 7.2 — Pairing.** Let $\text{embed} : \mathbb{N}^2 \rightarrow \mathbb{N}$ be the function defined as $\text{embed}(x_0, x_1) = \frac{1}{2}(x_0 + x_1)(x_0 + x_1 + 1) + x_1$.

1. Prove that for every $x^0, x^1 \in \mathbb{N}, \text{embed}(x^0, x^1)$ is indeed a natural number.

2. Prove that $\text{embed}$ is one-to-one.

3. Construct a NAND-TM program $P$ such that for every $x^0, x^1 \in \mathbb{N}$, $P(pf(x^0)pf(x^1)) = pf(\text{embed}(x^0, x^1))$, where $pf$ is the prefix-free encoding map defined above. You can use the syntactic sugar for inner loops, conditionals, and incrementing/decrementing the counter.
4. Construct NAND-TM programs $P_0, P_1$ such that for every $x^0, x^1 \in \mathbb{N}$ and $i \in \mathbb{N}$, $P_i(pf(embed(x^0, x^1))) = pf(x^i)$. You can use the syntactic sugar for inner loops, conditionals, and incrementing/decrementing the counter.

**Exercise 7.3 — Shortest Path.** Let $SHORTPATH : \{0, 1\}^* \rightarrow \{0, 1\}^*$ be the function that on input a string encoding a triple $(G, u, v)$ outputs a string encoding $\infty$ if $u$ and $v$ are disconnected in $G$ or a string encoding the length $k$ of the shortest path from $u$ to $v$. Prove that $SHORTPATH$ is computable by a Turing machine. See footnote for hint.\(^5\)

**Exercise 7.4 — Longest Path.** Let $LONGPATH : \{0, 1\}^* \rightarrow \{0, 1\}^*$ be the function that on input a string encoding a triple $(G, u, v)$ outputs a string encoding $\infty$ if $u$ and $v$ are disconnected in $G$ or a string encoding the length $k$ of the longest simple path from $u$ to $v$. Prove that $LONGPATH$ is computable by a Turing machine. See footnote for hint.\(^6\)

**Exercise 7.5 — Shortest path \(\lambda\) expression.** Let $SHORTPATH$ be as in Exercise 7.3. Prove that there exists a \(\lambda\) expression that computes $SHORTPATH$. You can use Exercise 7.3.

**Exercise 7.6 — Next-step function is local.** Prove Lemma 7.9 and use it to complete the proof of Theorem 7.7.

**Exercise 7.7 — \(\lambda\) calculus requires at most three variables.** Prove that for every \(\lambda\)-expression $e$ with no free variables there is an equivalent \(\lambda\)-expression $f$ that only uses the variables $x,y,$ and $z$.\(^7\)

**Exercise 7.8 — Evaluation order example in \(\lambda\) calculus.** 1. Let $e = \lambda x.7 ((\lambda x.xx)(\lambda x.xx))$. Prove that the simplification process of $e$ ends in a definite number if we use the “call by name” evaluation order while it never ends if we use the “call by value” order.

2. (bonus, challenging) Let $e$ be any \(\lambda\) expression. Prove that if the simplification process ends in a definite number if we use the “call by value” order then it also ends in such a number if we use the “call by name” order. See footnote for hint.\(^5\)

**Exercise 7.9 — Zip function.** Give an enhanced \(\lambda\) calculus expression to compute the function $zip$ that on input a pair of lists $I$ and $L$ of the

\(^5\) You don’t have to give a full description of a Turing machine: use our “eat the cake and have it too” paradigm to show the existence of such a machine by arguing from more powerful equivalent models.

\(^6\) Same hint as Exercise 7.4 applies. Note that for showing that $LONGPATH$ is computable you don’t have to give an efficient algorithm.

\(^7\) Hint: You can reduce the number of variables a function takes by “pairing them up”. That is, define a \(\lambda\) expression $PAIR$ such that for every $x, y$ $PAIRxy$ is some function $f$ such that $f0 = x$ and $f1 = y$. Then use $PAIR$ to iteratively reduce the number of variables used.

\(^8\) Use structural induction on the expression $e$. 
same length \( n \), outputs a list of \( n \) pairs \( M \) such that the \( j \)-th element of \( M \) (which we denote by \( M_j \)) is the pair \((I_j, L_j)\). Thus zip “zips together” these two lists of elements into a single list of pairs.\(^9\)

Exercise 7.10 — Next-step function without \textit{RECURSE}. Let \( M \) be a Turing machine. Give an enhanced \( \lambda \) calculus expression to compute the next-step function \( \text{NEXT}_M \) of \( M \) (as in the proof of Exercise 7.10) without using \textit{RECURSE}. See footnote for hint.\(^10\)

Exercise 7.11 — \( \lambda \) calculus to NAND-TM compiler (challenging). Give a program in the programming language of your choice that takes as input a \( \lambda \) expression \( e \) and outputs a NAND-TM program \( P \) that computes the same function as \( e \). For partial credit you can use the 6070 and all NAND-CIRC syntactic sugar in your output program. You can use any encoding of \( \lambda \) expressions as binary string that is convenient for you. See footnote for hint.\(^11\)

Exercise 7.12 — At least two in \( \lambda \) calculus. Let \( 1 = \lambda x, y. x \) and \( 0 = \lambda x, y. y \) as before. Define

\[
ALT = \lambda a, b, c. (a(b1(c10))(bc0))
\]  

(7.26)

Prove that \( ALT \) is a \( \lambda \) expression that computes the at least two function. That is, for every \( a, b, c \in \{0, 1\} \) (as encoded above) \( ALT abc = 1 \) if and only at least two of \( \{a, b, c\} \) are equal to 1.

Exercise 7.13 — Locality of next-step function. This question will help you get a better sense of the notion of \textit{locality of the next step function} of Turing Machines. This locality plays an important role in results such as the Turing completeness of \( \lambda \) calculus and one dimensional cellular automata, as well as results such as Godel’s Incompleteness Theorem and the Cook Levin theorem that we will see later in this course. Define \text{STRINGS} to be the a programming language that has the following semantics:

- A \text{STRINGS} program \( Q \) has a single string variable \texttt{str} that is both the input and the output of \( Q \). The program has no loops and no other variables, but rather consists of a sequence of conditional search and replace operations that modify \texttt{str}.

- The operations of a \text{STRINGS} program are:
  - REPLACE(pattern1, pattern2) where pattern1 and pattern2 are fixed strings. This replaces the first occurrence of pattern1 in \texttt{str} with pattern2

\(^9\) The name \textit{zip} is a common name for this operation, for example in Python. It should not be be confused with the \textit{zip} compression file format.

\(^10\) Use \textit{MAP} and \textit{REDUCE} (and potentially \textit{FILTER}). You might also find the function \textit{zip} of Exercise 7.9 useful.

\(^11\) Try to set up a procedure such that if array \texttt{Left} contains an encoding of a \( \lambda \) expression \( \lambda x. e \) and array \texttt{Right} contains an encoding of another \( \lambda \) expression \( e' \), then the array \texttt{Result} will contain \( e[x \rightarrow e'] \).
if search(pattern) { code } executes code if pattern is a substring of str. The code code can itself include nested if's. (One can also add an else { ... } to execute if pattern is not a substring of cond).

the returned value is str

A STRING program Q computes a function $F : \{0,1\}^* \rightarrow \{0,1\}^*$ if for every $x \in \{0,1\}^*$, if we initialize str to $x$ and then execute the sequence of instructions in $Q$, then at the end of the execution str equals $F(x)$.

For example, the following is a STRING program that computes the function $F : \{0,1\}^* \rightarrow \{0,1\}^*$ such that for every $x \in \{0,1\}^*$, if $x$ contains a substring of the form $y = 11ab11$ where $a, b \in \{0,1\}$, then $F(x) = x'$ where $x'$ is obtained by replacing the first occurrence of $y$ in $x$ with 00.

```c
if search('110011') {
    replace('110011','00')
} else if search('110111') {
    replace('110111','00')
} else if search('111011') {
    replace('111011','00')
} else if search('111111') {
    replace('111111','00')
}
```

Prove that for every Turing Machine program $M$, there exists a STRING program $Q$ that computes the $NEXT_M$ function that maps every string encoding a valid configuration of $M$ to the string encoding the configuration of the next step of $M$'s computation. (We don’t care what the function will do on strings that do not encode a valid configuration.) You don’t have to write the STRING program fully, but you do need to give a convincing argument that such a program exists.

### 7.10 BIBLIOGRAPHICAL NOTES

Chapters 7 in the wonderful book of Moore and Mertens [MM11] contains a great exposition much of this material. Chapter 3 in Savage’s book [Sav98] contains a more formal description of RAM machines, see also the paper [Hag98]. A study of RAM algorithms that are independent of the input size (known as the “transdichotomous RAM model”) was initiated by [FW93].

The RAM model can be very useful in studying the concrete complexity of practical algorithms. However, the exact set of operations
that are allowed in the RAM model at unit cost can vary between texts and contexts. One needs to be careful in making such definitions, especially if the word size grows, as was already shown by Shamir [Sha79].

The models of computation we considered so far are inherently sequential, but these days much computation happens in parallel, whether using multi-core processors or in massively parallel distributed computation in data centers or over the Internet. Parallel computing is important in practice, but it does not really make much difference for the question of what can and can’t be computed. After all, if a computation can be performed using $m$ machines in $t$ time, then it can be computed by a single machine in time $mt$.

The $\lambda$-calculus was described by Church in [Chu41]. Pierce’s book [Pie02] is a canonical textbook, see also [Bar84]. The “Currying technique” is named after the logician Haskell Curry (the Haskell programming language is named after Haskell Curry as well). Curry himself attributed this concept to Moses Schönfinkel, though for some reason the term “Schönfinkeling” never caught on.

Unlike most programming languages, the pure $\lambda$-calculus doesn’t have the notion of types. Every object in the $\lambda$ calculus can also be thought of as a $\lambda$ expression and hence as a function that takes one input and returns one output. All functions take one input and return one output, and if you feed a function an input of a form it didn’t expect, it still evaluates the $\lambda$ expression via “search and replace”, replacing all instances of its parameter with copies of the input expression you fed it. Typed variants of the $\lambda$ calculus are objects of intense research, and are strongly related to type systems for programming language and computer-verifiable proof systems, see [Pie02]. Some of the typed variants of the $\lambda$ calculus do not have infinite loops, which makes them very useful as ways of enabling static analysis of programs as well as computer-verifiable proofs. We will come back to this point in Chapter 9 and Chapter 21.

Tao has proposed showing the Turing completeness of fluid dynamics (a “water computer”) as a way of settling the question of the behavior of the Navier-Stokes equations, see this popular article.
Universality and uncomputability

“A function of a variable quantity is an analytic expression composed in any way whatsoever of the variable quantity and numbers or constant quantities.”, Leonhard Euler, 1748.

“The importance of the universal machine is clear. We do not need to have an infinity of different machines doing different jobs. … The engineering problem of producing various machines for various jobs is replaced by the office work of ‘programming’ the universal machine”, Alan Turing, 1948

One of the most significant results we showed for Boolean circuits (or equivalently, straight-line programs) is the notion of universality: there is a single circuit that can evaluate all other circuits. However, this result came with a significant caveat. To evaluate a circuit of $s$ gates, the universal circuit needed to use a number of gates larger than $s$. It turns out that uniform models such as Turing machines or NAND-TM programs allow us to “break out of this cycle” and obtain a truly universal Turing machine $U$ that can evaluate all other machines, including machines that are more complex (e.g., more states) than $U$ itself. (Similarly, there is a Universal NAND-TM program $U'$ that can evaluate all NAND-TM programs, including programs that have more lines than $U'$.)

It is no exaggeration to say that the existence of such a universal program/machine underlies the information technology revolution that began in the latter half of the 20th century (and is still ongoing). Up to that point in history, people have produced various special-purpose calculating devices such as the abacus, the slide ruler, and machines that compute various trigonometric series. But as Turing (who was perhaps the one to see most clearly the ramifications of universality) observed, a general purpose computer is much more powerful. Once build a device that can compute the single universal function we
have the ability, *via software*, to extend it to do arbitrary computations. For example, if we want to simulate a new Turing machine $M$, we do not need to build a new physical machine, but rather can represent $M$ as a string (i.e., using code) and then input $M$ to the universal machine $U$.

Beyond the practical applications, the existence of a universal algorithm also surprising theoretical ramifications, and in particular can be used to show the existence of *uncomputable functions*, upending the intuitions of mathematicians over the centuries from Euler to Hilbert. In this chapter we will prove the existence of the universal program, and also show its implications for uncomputability, see Fig. 8.1

8.1 UNIVERSALITY OR A META-CIRCULAR EVALUATOR

We start by proving the existence of a *universal Turing machine*. This is a single Turing machine $U$ that can evaluate arbitrary Turing machines $M$ on arbitrary inputs $x$, including machines $M$ that can have more states and larger alphabet than $U$ itself. In particular, $U$ can even be used to evaluate itself! This notion of self reference will appear time and again in this course, and as we will see, leads to several counter-intuitive phenomena in computing.

**Theorem 8.1 — Universal Turing Machine.** There exists a Turing machine $U$ such that on every string $M$ which represents a Turing machine, and $x \in \{0, 1\}^*$, $U(M, x) = M(x)$.

That is, if the machine $M$ halts on $x$ and outputs some $y \in \{0, 1\}^*$ then $U(M, x) = y$, and if $M$ does not halt on $x$ (i.e., $M(x) = \perp$) then $U(M, x) = \perp$.

**Figure 8.1:** In this chapter we will show the existence of a *universal Turing machine* and then use this to derive first the existence of some uncomputable function. We then use this to derive the uncomputability of Turing’s famous “halting problem” (i.e., the HALT function), from which we a host of other uncomputability results follow. We also introduce reductions, which allow us to use the uncomputability of a function $F$ to derive the uncomputability of a new function $G$. The cartoon of the Halting problem is copyright 2019 Charles F. Cooper.

8.2 REDUCTIONS

In this section we introduce reductions, which allow us to convert the uncomputability of a function $F$ to the uncomputability of another function $G$. This is done by showing that if $F$ is uncomputable, then $G$ is also uncomputable via a reduction from $F$ to $G$. A reduction from $F$ to $G$ is a Turing machine that takes input $x$ and computes $G(x)$ using $F$. We will see several examples of reductions in this section, including a reduction from the Halting problem to Rice’s Theorem.

**Figure 8.2:** A *Universal Turing Machine* is a single Turing Machine $U$ that can evaluate, given input the (description as a string of) arbitrary Turing machine $M$ and input $x$, the output of $M$ on $x$. In contrast to the universal circuit depicted in Fig. 5.6, the machine $M$ can be much more complex (e.g., more states or tape alphabet symbols) than $U$. 
There is a single algorithm that can evaluate arbitrary algorithms on arbitrary inputs.

**Proof Idea:**

Once you understand what the theorem says, it is not that hard to prove. The desired program $U$ is an *interpreter* for Turing machines. That is, $U$ gets a representation of the machine $M$ (think of it as source code), and some input $x$, and needs to simulate the execution of $M$ on $x$.

Think of how you would code $U$ in your favorite programming language. First, you would need to decide on some representation scheme for $M$. For example, you can use an array or a dictionary to encode $M$’s transition function. Then you would use some data structure, such as a list, to store the contents of $M$’s tape. Now you can simulate $M$ step by step, updating the data structure as you go along. The interpreter will continue the simulation until the machine halts.

Once you do that, translating this interpreter from your favorite programming language to a Turing machine can be done just as we have seen in Chapter 7. The end result is what’s known as a “meta-circular evaluator”: an interpreter for a programming language in the same one. This is a concept that has a long history in computer science starting from the original universal Turing machine. See also Fig. 8.3.

* 

### 8.1.1 Proving the existence of a universal Turing Machine

To prove (and even properly state) Theorem 8.1, we need fix some representation for Turing machines as strings. For example, one potential choice for such a representation is to use the equivalence between Turing machines and NAND-TM programs and hence represent a Turing machine $M$ using the ASCII encoding of the source code of the corresponding NAND-TM program $P$. However, we will use a more direct encoding.

Let $M$ be a Turing machine with $k$ states and a size $\ell$ alphabet $\Sigma = \{\sigma_0, \ldots, \sigma_{\ell-1}\}$ (we use the convention $\sigma_0 = 0, \sigma_1 = 1, \sigma_2 = \varnothing, \sigma_3 = \triangleright$). We represent $M$ by the triple $(k, \ell, T)$ where $T$ is the table of values for $\delta_M$:

$$T = (\delta_M(0, 0), \delta_M(0, \sigma_0), \ldots, \delta_M(k - 1, \sigma_{k-1})) , \quad (8.1)$$

where each value $\delta_M(s, \sigma)$ is a triple $(s', \sigma', d)$ with $s' \in [k], \sigma' \in \Sigma$ and $d$ a number $\{0, 1, 2, 3\}$ encoding one of $\{L, R, S, H\}$. Thus such a machine $M$ is encoded by a list of $2 + 3k \cdot \ell$ natural numbers, which can be represented as a binary string.

Using this representation, we can formally prove Theorem 8.1.
**Proof of Theorem 8.1.** We will only sketch the proof, giving the major ideas. First, we observe that we can easily write a Python program that, on input a representation \((k, \ell, T)\) of a Turing machine \(M\) and an input \(x\), evaluates \(M\) on \(X\). Here is the code of this program for concreteness, though you can feel free to skip it if you are not familiar with (or interested in) Python:

```python
# constants
Φ = 2  # empty symbol
Δ = 3  # starting symbol
L = 0  # go left
R = 1  # go right
S = 2  # stay
H = 3  # halt

def EVAL(k, l, T, x):
    """Evaluate a Turing machine on input x
    Turing machine is represented by:
    k: number of states
    l: number of symbols in alphabet (containing {0,1,Φ,Δ})
    T: transition function: a dictionary such that T[(s,a)] is equal to (_s,_a,_d)
    where s is old state, a is symbol read
    _s is new state
    _a is symbol to write
    _d in {L,R,S,H} is head movement
    """

    Tape = [ Δ ] # List/array containing contents of tape

    # Initialize with input
    for i in range(len(x)): Tape.append(int(x[i]))

    i = 0  # current position
    s = 0  # current state

    while True:
        a = Tape[i]  # read symbol
        _s, _a, _d = T[(s,a)]  # lookup transition table
        Tape[i] = _a  # write symbol
        s = _s  # update state
```

# move head:
if _d == H: break
if _d == L: i = max(i-1, 0)
if _d == R: i += 1

# add empty symbols to tape if needed
if i >= len(Tape): Tape.append(Φ)

# Scan tape for output
Y = []
j = 1
while j < len(Tape) and Tape[j] != Φ:
    Y.append(Tape[j])
j += 1

return Y

The above does not prove the theorem as stated, since we need to show a Turing machine that computes EVAL rather than a Python program. With enough effort, we can translate this Python code line by line to a Turing machine. However, to prove the theorem we don’t need to do this, but can use our “eat the cake and have it too” paradigm. That is, while we need to evaluate a Turing machine, in writing the code for the interpreter we are allowed to use a richer model such as NAND-RAM since it is equivalent in power to Turing machines per Theorem 7.1).

Translating the above Python code to NAND-RAM is truly straightforward. The only issue is that NAND-RAM doesn’t have the dictionary data structure built in, which we have used above to store the transition function T. However, we can represent a dictionary D of the form \{key0 : val0, ... , keym-1 : val_{m-1}\} as simply a list of pairs. To compute D[k] we can scan over all the pairs until we find one of the form (k, v) in which case we return v. Similarly we scan the list to update the dictionary with a new value, either modifying it or appending the pair (key, val) at the end.

Remark 8.2 — Efficiency of the simulation. The argument in the proof of Theorem 8.1 is a very inefficient way to implement the dictionary data structure in practice, but it suffices for the purpose of proving the theorem. Reading and writing to a dictionary of m values in this implementation takes \Omega(m) steps, but
it is in fact possible to do this in $O(\log m)$ steps using a search tree data structure or even $O(1)$ (for “typical” instances) using a hash table. NAND-RAM and RAM machines correspond to the architecture of modern electronic computers, and so we can implement hash tables and search trees in NAND-RAM just as they are implemented in other programming languages.

8.1.2 Implications of universality (discussion)

Figure 8.3: a) A particularly elegant example of a “meta-circular evaluator” comes from John McCarthy’s 1960 paper, where he defined the Lisp programming language and gave a Lisp function that evaluates an arbitrary Lisp program (see above). Lisp was not initially intended as a practical programming language and this example was merely meant as an illustration that the Lisp universal function is more elegant than the universal Turing machine. It was McCarthy’s graduate student Steve Russell who suggested that it can be implemented. As McCarthy later recalled, “I said to him, ho, ho, you’re confusing theory with practice, this eval is intended for reading, not for computing. But he went ahead and did it. That is, he compiled the eval in my paper into IBM 704 machine code, fixing a bug, and then advertised this as a Lisp interpreter, which it certainly was”. b) A self-replicating C program from the classic essay of Thompson [Tho84].

There is more than one Turing machine $U$ that satisfies the conditions of Theorem 8.1, but the existence of even a single such machine is already extremely fundamental to both the theory and practice of computer science. Theorem 8.1’s impact reaches beyond the particular model of Turing machines. Because we can simulate every Turing Machine by a NAND-TM program and vice versa, Theorem 8.1 immediately implies there exists a universal NAND-TM program $P_U$ such that $P_U(P, x) = P(x)$ for every NAND-TM program $P$. We can also “mix and match” models. For example since we can simulate every NAND-RAM program by a Turing machine, and every Turing Machine by the $\lambda$ calculus, Theorem 8.1 implies that there exists a $\lambda$ expression $e$ such that for every NAND-RAM program $P$ and input $x$ on which $P(x) = y$, if we encode $(P, x)$ as a $\lambda$-expression $f$ (using the $\lambda$-calculus encoding of strings as lists of 0’s and 1’s) then $(e \, f)$ evaluates to an encoding of $y$. More generally we can say that for every $\mathcal{X}$ and $\mathcal{Y}$ in the set \{ Turing Machines, RAM Machines, NAND-TM, NAND-RAM, $\lambda$-calculus, JavaScript, Python, … \} of Turing equivalent models, there exists a program/machine in $\mathcal{X}$ that computes the map $(P, x) \mapsto P(x)$ for every program/machine $P \in \mathcal{Y}$. 
The idea of a “universal program” is of course not limited to theory. For example, compilers for programming languages are often used to compile themselves, as well as programs more complicated than the compiler. (An extreme example of this is Fabrice Bellard’s Obfuscated Tiny C Compiler which is a C program of 2048 bytes that can compile a large subset of the C programming language, and in particular can compile itself.) This is also related to the fact that it is possible to write a program that can print its own source code, see Fig. 8.3.

8.2 IS EVERY FUNCTION COMPUTABLE?

In Theorem 4.12, we saw that NAND-CIRC programs can compute every finite function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \). Therefore a natural guess is that NAND-TM programs (or equivalently, Turing Machines) could compute every infinite function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \). However, this turns out to be false. That is, there exists a function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) that is uncomputable!

The existence of uncomputable functions is quite surprising. Our intuitive notion of a “function” (and the notion most mathematicians had until the 20th century) is that a function \( f \) defines some implicit or explicit way of computing the output \( f(x) \) from the input \( x \). The notion of an “uncomputable function” thus seems to be a contradiction in terms, but yet the following theorem shows that such creatures do exist:

**Theorem 8.3 — Uncomputable functions.** There exists a function \( F^* : \{0, 1\}^* \rightarrow \{0, 1\} \) that is not computable by any Turing machine.

**Proof Idea:**

The idea behind the proof follows quite closely Cantor’s proof that the reals are uncountable (Theorem 2.5), and in fact the theorem can also be obtained fairly directly from that result (see Exercise 6.11). However, it is instructive to see the direct proof. The idea is to construct \( F^* \) in a way that will ensure that every possible machine \( M \) will in fact fail to compute \( F^* \). We do so by defining \( F^*(x) \) to equal 1 if \( x \) describes a Turing machine \( M \) which satisfies \( M(x) = 1 \) and defining \( F^*(x) = 0 \) otherwise. By construction, if \( M \) is any Turing machine and \( x \) is the string describing it, then \( F^*(x) \neq M(x) \) and therefore \( M \) does not compute \( F^* \).

\[
F^*(x) = \begin{cases} 1 & \text{if } x \text{ satisfies (1)} \text{,}\, M(x) = 1 \\ 0 & \text{otherwise} \end{cases}
\]

*Proof of Theorem 8.3.* The proof is illustrated in Fig. 8.4. We start by defining the following function \( G : \{0, 1\}^* \rightarrow \{0, 1\} \):

For every string \( x \in \{0, 1\}^* \), if \( x \) satisfies (1) \( x \) is a valid representation of some Turing machine \( M \) (per the representation scheme
above) and (2) when the program $M$ is executed on the input $x$ it halts and produces an output, then we define $G(x)$ as the first bit of this output. Otherwise (i.e., if $x$ is not a valid representation of a Turing machine, or the machine $M_x$ never halts on $x$) we define $G(x) = 0$. We define $F^*(x) = 1 - G(x)$.

We claim that there is no Turing machine that computes $F^*$. Indeed, suppose, towards the sake of contradiction, that exists a machine $M$ that computes $F^*$, and let $x$ be the binary string that represents the machine $M$. On one hand, since by our assumption $M$ computes $F^*$, on input $x$ the machine $M$ halts and outputs $F^*(x)$. On the other hand, by the definition of $F^*$, since $x$ is the representation of the machine $M$, $F^*(x) = 1 - G(x) = 1 - M(x)$, hence yielding a contradiction.

The proof of Theorem 8.3 is short but subtle. I suggest that you pause here and go back to read it again and think about it - this is a proof that is worth reading at least twice if not three or four times. It is not often the case that a few lines of mathematical reasoning establish a deeply profound fact - that there are problems we simply cannot solve.

The type of argument used to prove Theorem 8.3 is known as diagonalization since it can be described as defining a function based on the diagonal entries of a table as in Fig. 8.4. The proof can be thought of as an infinite version of the counting argument we used for showing lower bound for NAND-CIRC programs in Theorem 5.3.
Namely, we show that it’s not possible to compute all functions from \( \{0, 1\}^* \rightarrow \{0, 1\} \) by Turing machines simply because there are more functions like that than there are Turing machines.

As mentioned in ??, many texts use the “language” terminology and so will call a set \( L \subseteq \{0, 1\}^* \) an **undecidable** or **non recursive language** if the function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) such that \( F(x) = 1 \leftrightarrow x \in L \) is uncomputable.

### 8.3 THE HALTING PROBLEM

**Theorem 8.3** shows that there is *some* function that cannot be computed. But is this function the equivalent of the “tree that falls in the forest with no one hearing it”? That is, perhaps it is a function that no one actually wants to compute. It turns out that there are natural uncomputable functions:

**Theorem 4 — Uncomputability of Halting function.** Let \( HALT : \{0, 1\}^* \rightarrow \{0, 1\} \) be the function such that for every string \( M \in \{0, 1\}^* \),

\[
HALT(M, x) = 1 \quad \text{if } \text{Turing machine } M \text{ halts on the input } x \quad \text{and} \quad HALT(M, x) = 0 \quad \text{otherwise. Then } HALT \text{ is not computable.}
\]

Before turning to prove **Theorem 8.4**, we note that \( HALT \) is a very natural function to want to compute. For example, one can think of \( HALT \) as a special case of the task of managing an “App store”. That is, given the code of some application, the gatekeeper for the store needs to decide if this code is safe enough to allow in the store or not. At a minimum, it seems that we should verify that the code would not go into an infinite loop.

**Proof Idea:**

One way to think about this proof is as follows:

\[
\text{Uncomputability of } F^* + \text{ Universality} = \text{ Uncomputability of } HALT
\]

That is, we will use the universal Turing machine that computes \( EVAL \) to derive the uncomputability of \( HALT \) from the uncomputability of \( F^* \) shown in **Theorem 8.3**. Specifically, the proof will be by contradiction. That is, we will assume towards a contradiction that \( HALT \) is computable, and use that assumption, together with the universal Turing machine of **Theorem 8.1**, to derive that \( F^* \) is computable, which will contradict **Theorem 8.3**.

*
**Big Idea 11** If a function $F$ is uncomputable we can show that another function $H$ is uncomputable by giving a way to *reduce* the task of computing $F$ to computing $H$.

*Proof of Theorem 8.4.* The proof will use the previously established result Theorem 8.3. Recall that Theorem 8.3 shows that the following function $F^*: \{0, 1\}^* \to \{0, 1\}$ is uncomputable:

$$F^*(x) = \begin{cases} 1 & x(x) = 0 \\ 0 & \text{otherwise} \end{cases} \quad (8.3)$$

where $x(x)$ denotes the output of the Turing machine described by the string $x$ on the input $x$ (with the usual convention that $x(x) = \bot$ if this computation does not halt).

We will show that the uncomputability of $F^*$ implies the uncomputability of $HALT$. Specifically, we will assume, towards a contradiction, that there exists a Turing machine $M$ that can compute the $HALT$ function, and use that to obtain a Turing machine $M'$ that computes the function $F^*$. (This is known as a proof by *reduction*, since we reduce the task of computing $F^*$ to the task of computing $HALT$. By the contrapositive, this means the uncomputability of $F^*$ implies the uncomputability of $HALT$.)

Indeed, suppose that $M$ is a Turing machine that computes $HALT$. Algorithm 8.5 describes a Turing Machine $M'$ that computes $F^*$. (We use “high level” description of Turing machines, appealing to the “have your cake and eat it too” paradigm, see Big Idea 9.)

```
Algorithm 8.5 — $F^*$ to $HALT$ reduction.

Input: $x \in \{0, 1\}^*$
Output: $F^*(x)$

1: # Assume T.M. $M_{HALT}$ computes $HALT$
2: Let $z \leftarrow M_{HALT}(x, x)$. # Assume $z = HALT(x, x)$.
3: if $z = 0$ then
4:    return 0
5: end if
6: Let $y \leftarrow U(x, x)$ # $U$ universal TM, i.e., $y = x(x)$
7: if $y = 0$ then
8:    return 1
9: end if
10: return 0
```

We claim that Algorithm 8.5 computes the function $F^*$. Indeed, suppose that $x(x) = 0$ (and hence $F^*(x) = 1$). In this
case, $\text{HALT}(x, x) = 1$ and hence, under our assumption that $M(x, x) = \text{HALT}(x, x)$, the value $z$ will equal 1, and hence Algorithm 8.5 will set $y = x(x) = 0$, and output the correct value 1.

Suppose otherwise that $x(x) \neq 0$ (and hence $F^*(x) = 0$). In this case there are two possibilities:

- **Case 1**: The machine described by $x$ does not halt on the input $x$. In this case, $\text{HALT}(x, x) = 0$. Since we assume that $M$ computes $\text{HALT}$ it means that on input $x, x$, the machine $M$ must halt and output the value 0. This means that Algorithm 8.5 will set $z = 0$ and output 0.

- **Case 2**: The machine described by $x$ halts on the input $x$ and outputs some $y' \neq 0$. In this case, since $\text{HALT}(x, x) = 1$, under our assumptions, Algorithm 8.5 will set $y = y' \neq 0$ and so output 0.

We see that in all cases, $M'(x) = F^*(x)$, which contradicts the fact that $F^*$ is uncomputable. Hence we reach a contradiction to our original assumption that $M$ computes $\text{HALT}$.

---

8.3.1 *Is the Halting problem really hard? (discussion)*

Many people’s first instinct when they see the proof of Theorem 8.4 is to not believe it. That is, most people do believe the mathematical statement, but intuitively it doesn’t seem that the Halting problem is really that hard. After all, being uncomputable only means that $\text{HALT}$ cannot be computed by a Turing machine.

But programmers seem to solve $\text{HALT}$ all the time by informally or formally arguing that their programs halt. It’s true that their programs are written in C or Python, as opposed to Turing machines, but that makes no difference: we can easily translate back and forth between this model and any other programming language.

While every programmer encounters at some point an infinite loop, is there really no way to solve the halting problem? Some people
argue that they personally can, if they think hard enough, determine whether any concrete program that they are given will halt or not. Some have even argued that humans in general have the ability to do that, and hence humans have inherently superior intelligence to computers or anything else modeled by Turing machines.¹

The best answer we have so far is that there truly is no way to solve \textsc{HALT}, whether using Macs, PCs, quantum computers, humans, or any other combination of electronic, mechanical, and biological devices. Indeed this assertion is the content of the Church-Turing Thesis. This of course does not mean that for every possible program \( P \), it is hard to decide if \( P \) enters an infinite loop. Some programs don’t even have loops at all (and hence trivially halt), and there are many other far less trivial examples of programs that we can certify to never enter an infinite loop (or programs that we know for sure that \textit{will} enter such a loop). However, there is no general procedure that would determine for an arbitrary program \( P \) whether it halts or not. Moreover, there are some very simple programs for which no one knows whether they halt or not. For example, the following Python program will halt if and only if Goldbach’s conjecture is false:

```python
def isprime(p):
    return all(p % i for i in range(2, p - 1))

def Goldbach(n):
    return any(isprime(p) and isprime(n - p) for p in range(2, n - 1))

n = 4
while True:
    if not Goldbach(n):
        break
    n += 2
```

Given that Goldbach’s Conjecture has been open since 1742, it is unclear that humans have any magical ability to say whether this (or other similar programs) will halt or not.

8.3.2 A direct proof of the uncomputability of \textsc{HALT} (optional)
It turns out that we can combine the ideas of the proofs of Theorem 8.3 and Theorem 8.4 to obtain a short proof of the latter theorem, that does not appeal to the uncomputability of \( F^* \). This short proof appeared in print in a 1965 letter to the editor of Christopher Strachey:

```
To the Editor, The Computer Journal.
An Impossible Program
Sir,
```

Figure 8.5: SMBC’s take on solving the Halting problem.
A well-known piece of folk-lore among programmers holds that it is impossible to write a program which can examine any other program and tell, in every case, if it will terminate or get into a closed loop when it is run. I have never actually seen a proof of this in print, and though Alan Turing once gave me a verbal proof (in a railway carriage on the way to a Conference at the NPL in 1953), I unfortunately and promptly forgot the details. This left me with an uneasy feeling that the proof must be long or complicated, but in fact it is so short and simple that it may be of interest to casual readers. The version below uses CPL, but not in any essential way.

Suppose $T[R]$ is a Boolean function taking a routine (or program) $R$ with no formal or free variables as its arguments and that for all $R$, $T[R] = True$ if $R$ terminates if run and that $T[R] = False$ if $R$ does not terminate.

Consider the routine $P$ defined as follows:

```plaintext
rec routine P
  $L$: if $T[P]$ go to $L$
  Return $\$

If $T[P] = True$ the routine $P$ will loop, and it will only terminate if $T[P] = False$. In each case $T[P]$" has exactly the wrong value, and this contradiction shows that the function $T$ cannot exist.

Yours faithfully,

C. Strachey

Churchill College, Cambridge
```

Try to stop and extract the argument for proving Theorem 8.4 from the letter above.

Since CPL is not as common today, let us reproduce this proof. The idea is the following: suppose for the sake of contradiction that there exists a program $T$ such that $T(f, x)$ equals True iff $f$ halts on input $x$. (Strachey’s letter considers the no-input variant of $HALT$, but as we’ll see, this is an immaterial distinction.) Then we can construct a program $P$ and an input $x$ such that $T(P, x)$ gives the wrong answer. The idea is that on input $x$, the program $P$ will do the following: run $T(x, x)$, and if the answer is True then go into an infinite loop, and otherwise halt. Now you can see that $T(P, P)$ will give the wrong answer: if $P$ halts when it gets its own code as input, then $T(P, P)$ is
supposed to be True, but then \( P(P) \) will go into an infinite loop. And if \( P \) does not halt, then \( T(P,P) \) is supposed to be False but then \( P(P) \) will halt. We can also code this up in Python:

```python
def CantSolveMe(T):
    """
    Gets function \( T \) that claims to solve HALT.
    Returns a pair \((P,x)\) of code and input on which
    \( T(P,x) \neq \text{HALT}(x) \)
    """
    def fool(x):
        if T(x,x):
            while True: pass
        return "I halted"
    return (fool,fool)
```

For example, consider the following Naive Python program \( T \) that guesses that a given function does not halt if its input contains `while` or `for`:

```python
def T(f,x):
    """
    Crude halting tester - decides it doesn't halt if it contains a loop.
    """
    import inspect
    source = inspect.getsource(f)
    if source.find("while"): return False
    if source.find("for"): return False
    return True
```

If we now set \((f,x) = \text{CantSolveMe}(T)\), then \( T(f,x) = \text{False} \) but \( f(x) \) does in fact halt. This is of course not specific to this particular \( T \): for every program \( T \), if we run \((f,x) = \text{CantSolveMe}(T)\) then we’ll get an input on which \( T \) gives the wrong answer to \( \text{HALT} \).

### 8.4 Reductions

The Halting problem turns out to be a linchpin of uncomputability, in the sense that Theorem 8.4 has been used to show the uncomputability of a great many interesting functions. We will see several examples in such results in this chapter and the exercises, but there are many more such results (see Fig. 8.6).

The idea behind such uncomputability results is conceptually simple but can at first be quite confusing. If we know that \( \text{HALT} \) is uncomputable, and we want to show that some other function \( \text{BLAH} \) is uncomputable, then we can do so via a contrapositive argument (i.e.,
proof by contradiction). That is, we show that if there exists Turing machine that computes \textit{BLAH} then there exists a Turing machine that computes \textit{HALT}. (Indeed, this exactly how we showed that \textit{HALT} itself is uncomputable, by reducing this fact to the uncomputability of the function \textit{F} from Theorem 8.3.)

For example, to prove that \textit{BLAH} is uncomputable, we could show that there is a computable function \( R : \{0, 1\}^* \rightarrow \{0, 1\}^* \) such that for every pair \( M \) and \( x \), \( \text{HALT}(M, x) = \text{BLAH}(R(M, x)) \). The existence of such a function \( R \) implies that if \textit{BLAH} was computable then \textit{HALT} would be computable as well, hence leading to a contradiction! The confusing part about reductions is that we are assuming something we believe is false (that \textit{BLAH} has an algorithm) to derive something that we know is false (that \textit{HALT} has an algorithm). Michael Sipser describes such results as having the form “If pigs could whistle then horses could fly”.

A reduction-based proof has two components. For starters, since we need \( R \) to be computable, we should describe the algorithm to compute it. The algorithm to compute \( R \) is known as a reduction since the transformation \( R \) modifies an input to \textit{HALT} to an input to \textit{BLAH}, and hence reduces the task of computing \textit{HALT} to the task of computing \textit{BLAH}. The second component of a reduction-based proof is the analysis of the algorithm \( R \): namely a proof that \( R \) does indeed satisfy the desired properties.

Reduction-based proofs are just like other proofs by contradiction, but the fact that they involve hypothetical algorithms that don’t really exist tends to make reductions quite confusing. The one silver lining is that at the end of the day the notion of reductions is mathematically quite simple, and so it’s not that bad even if you have to go back to
first principles every time you need to remember what is the direction that a reduction should go in.

**Remark 8.6 — Reductions are algorithms.** A reduction is an algorithm, which means that, as discussed in Remark 0.3, a reduction has three components:

- **Specification (what):** In the case of a reduction from HALT to BLAH, the specification is that function $R: \{0, 1\}^* \rightarrow \{0, 1\}^*$ should satisfy that $HALT(M, x) = BLAH(R(M, x))$ for every Turing machine $M$ and input $x$. In general, to reduce a function $F$ to $G$, the reduction should satisfy $F(w) = G(R(w))$ for every input $w$.

- **Implementation (how):** The algorithm’s description: the precise instructions how to transform an input $w$ to the output $R(w)$.

- **Analysis (why):** A proof that the algorithm meets the specification. In particular, in a reduction from $F$ to $G$ this is a proof that for every input $w$, the output $y$ of the algorithm satisfies that $F(w) = G(y)$.

### 8.4.1 Example: Halting on the zero problem

Here is a concrete example for a proof by reduction. We define the function $HALTONZERO: \{0, 1\}^* \rightarrow \{0, 1\}$ as follows. Given any string $M$, $HALTONZERO(M) = 1$ if and only if $M$ describes a Turing machine that halts when it is given the string $0$ as input. A priori $HALTONZERO$ seems like a potentially easier function to compute than the full-fledged $HALT$ function, and so we could perhaps hope that it is not uncomputable. Alas, the following theorem shows that this is not the case:

**Theorem 8.7 — Halting without input.** $HALTONZERO$ is uncomputable.

The proof of Theorem 8.7 is below, but before reading it you might want to pause for a couple of minutes and think how you would prove it yourself. In particular, try to think of what a reduction from $HALT$ to $HALTONZERO$ would look like. Doing so is an excellent way to get some initial comfort with the notion of proofs by reduction, which a technique we will be using time and again in this book.
Fig. 8.7: To prove Theorem 8.7, we show that HALTONZERO is uncomputable by giving a reduction from the task of computing HALT to the task of computing HALTONZERO. This shows that if there was a hypothetical algorithm $A$ computing HALTONZERO, then there would be an algorithm $B$ computing HALT, contradicting Theorem 8.4. Since neither $A$ nor $B$ actually exists, this is an example of an implication of the form “if pigs could whistle then horses could fly”.

Proof of Theorem 8.7. The proof is by reduction from HALT, see Fig. 8.7. We will assume, towards the sake of contradiction, that HALTONZERO is computable by some algorithm $A$, and use this hypothetical algorithm $A$ to construct an algorithm $B$ to compute HALT, hence obtaining a contradiction to Theorem 8.4. (As discussed in ??, following our “eat your cake and have it too” paradigm, we just use the generic name “algorithm” rather than worrying whether we model them as Turing machines, NAND-TM programs, NAND-RAM, etc.; this makes no difference since all these models are equivalent to one another.)

Since this is our first proof by reduction from the Halting problem, we will spell it out in more details than usual. Such a proof by reduction consists of two steps:

1. **Description of the reduction:** We will describe the operation of our algorithm $B$, and how it makes “function calls” to the hypothetical algorithm $A$.

2. **Analysis of the reduction:** We will then prove that under the hypothesis that Algorithm $A$ computes HALTONZERO, Algorithm $B$ will compute HALT.
Our Algorithm $B$ works as follows: on input $M, x$, it runs Algorithm 8.8 to obtain a Turing Machine $M'$, and then returns $A(M')$. The machine $M'$ ignores its input $z$ and simply runs $M$ on $x$.

In pseudocode, the program $N_{M,x}$ will look something like the following:

```python
def N(z):
    M = r'.......'
    # a string constant containing desc. of M
    x = r'.......'
    # a string constant containing x
    return eval(M, x)
    # note that we ignore the input z
```

That is, if we think of $N_{M,x}$ as a program, then it is a program that contains $M$ and $x$ as “hardwired constants”, and given any input $z$, it simply ignores the input and always returns the result of evaluating $M$ on $x$. The algorithm $B$ does not actually execute the machine $N_{M,x}$. $B$ merely writes down the description of $N_{M,x}$ as a string (just as we did above) and feeds this string as input to $A$.

The above completes the description of the reduction. The analysis is obtained by proving the following claim:

**Claim:** For every strings $M, x, z$, the machine $N_{M,x}$ constructed by Algorithm $B$ in Step 1 satisfies that $N_{M,x}$ halts on $z$ if and only if the program described by $M$ halts on the input $x$.

**Proof of Claim:** Since $N_{M,x}$ ignores its input and evaluates $M$ on $x$ using the universal Turing machine, it will halt on $z$ if and only if $M$ halts on $x$.

In particular if we instantiate this claim with the input $z = 0$ to $N_{M,x}$, we see that $\text{HALTONZERO}(N_{M,x}) = \text{HALT}(M, x)$. Thus if the hypothetical algorithm $A$ satisfies $A(M) = \text{HALTONZERO}(M)$ for every $M$ then the algorithm $B$ we construct satisfies $B(M, x) =
HALT(M, x) for every M, x, contradicting the uncomputability of HALT.

**Remark 8.9 — The hardwiring technique.** In the proof of Theorem 8.7 we used the technique of “hardwiring” an input x to a program/machine P. That is, modifying a program P that it uses “hardwired constants” for some of all of its input. This technique is quite common in reductions and elsewhere, and we will often use it again in this course.

### 8.5 Rice’s Theorem and the Impossibility of General Software Verification

The uncomputability of the Halting problem turns out to be a special case of a much more general phenomenon. Namely, that we cannot certify semantic properties of general purpose programs. “Semantic properties” mean properties of the function that the program computes, as opposed to properties that depend on the particular syntax used by the program.

An example for a semantic property of a program P is the property that whenever P is given an input string with an even number of 1’s, it outputs 0. Another example is the property that P will always halt whenever the input ends with a 1. In contrast, the property that a C program contains a comment before every function declaration is not a semantic property, since it depends on the actual source code as opposed to the input/output relation.

Checking semantic properties of programs is of great interest, as it corresponds to checking whether a program conforms to a specification. Alas it turns out that such properties are in general uncomputable. We have already seen some examples of uncomputable semantic functions, namely HALT and HALTONZERO, but these are just the “tip of the iceberg”. We start by observing one more such example:

**Theorem 8.10 — Computing all zero function.** Let ZEROFUNC : \{0, 1\}∗ → \{0, 1\} be the function such that for every M ∈ \{0, 1\}∗, ZEROFUNC(M) = 1 if and only if M represents a Turing machine such that M outputs 0 on every input x ∈ \{0, 1\}∗. Then ZEROFUNC is uncomputable.

Despite the similarity in their names, ZEROFUNC and HALTONZERO are two different functions. For exam-
ple, if $M$ is a Turing machine that on input $x \in \{0, 1\}^*$, halts and outputs the OR of all of $x$’s coordinates, then $\text{HALTONZERO}(M) = 1$ (since $M$ does halt on the input $0$) but $\text{ZEROFUNC}(M) = 0$ (since $M$ does not compute the constant zero function).

**Proof of Theorem 8.10.** The proof is by reduction to $\text{HALTONZERO}$. Suppose, towards the sake of contradiction, that there was an algorithm $A$ such that $A(M) = \text{ZEROFUNC}(M)$ for every $M \in \{0, 1\}^*$. Then we will construct an algorithm $B$ that solves $\text{HALTONZERO}$, contradicting Theorem 8.7.

Given a Turing machine $N$ (which is the input to $\text{HALTONZERO}$), our Algorithm $B$ does the following:

1. Construct a Turing Machine $M$ which on input $x \in \{0, 1\}^*$, first runs $N(0)$ and then outputs $0$.
2. Return $A(M)$.

Now if $N$ halts on the input $0$ then the Turing machine $M$ computes the constant zero function, and hence under our assumption that $A$ computes $\text{ZEROFUNC}$, $A(M) = 1$. If $N$ does not halt on the input $0$, then the Turing machine $M$ will not halt on any input, and so in particular will not compute the constant zero function. Hence under our assumption that $A$ computes $\text{ZEROFUNC}$, $A(M) = 0$. We see that in both cases, $\text{ZEROFUNC}(M) = \text{HALTONZERO}(N)$ and hence the value that Algorithm $B$ returns in step 2 is equal to $\text{HALTONZERO}(N)$ which is what we needed to prove.

Another result along similar lines is the following:

**Theorem 8.11 — Uncomputability of verifying parity.** The following function is uncomputable

$$
\text{COMPUTES-PARITY}(P) = \begin{cases} 
1 & P \text{ computes the parity function} \\
0 & \text{otherwise}
\end{cases}
$$

(8.4)

We leave the proof of Theorem 8.11 as an exercise (Exercise 8.6). I strongly encourage you to stop here and try to solve this exercise.
8.5.1 Rice’s Theorem

Theorem 8.11 can be generalized far beyond the parity function. In fact, this generalization rules out verifying any type of semantic specification on programs. We define a semantic specification on programs to be some property that does not depend on the code of the program but just on the function that the program computes.

For example, consider the following two C programs

```c
int First(int k) {
    return 2*k;
}
int Second(int n) {
    int i = 0;
    int j = 0
    while (j<n) {
        i = i + 2;
        j = j + 1;
    }
    return i;
}
```

First and Second are two distinct C programs, but they compute the same function. A semantic property, would be either true for both programs or false for both programs, since it depends on the function the programs compute and not on their code. An example for a semantic property that both First and Second satisfy is the following: “The program \( P \) computes a function \( f \) mapping integers to integers satisfying that \( f(n) \geq n \) for every input \( n \)”.

A property is not semantic if it depends on the source code rather than the input/output behavior. For example, properties such as “the program contains the variable \( k \)” or “the program uses the while operation” are not semantic. Such properties can be true for one of the programs and false for the others. Formally, we define semantic properties as follows:

**Definition 8.12 — Semantic properties.** A pair of Turing machines \( M \) and \( M' \) are functionally equivalent if for every \( x \in \{0, 1\}^* \), \( M(x) = M'(x) \). (In particular, \( M(x) = \bot \) iff \( M'(x) = \bot \) for all \( x \).)

A function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) is semantic if for every pair of strings \( M, M' \) that represent functionally equivalent Turing machines, \( F(M) = F(M') \).

There are two trivial examples of semantic functions: the constant one function and the constant zero function. For example, if \( Z \) is the
constant zero function (i.e., \( Z(M) = 0 \) for every \( M \)) then clearly \( F(M) = F(M') \) for every pair of Turing machines \( M \) and \( M' \) that are functionally equivalent \( M \) and \( M' \). Here is a non-trivial example

**Solved Exercise 8.1 — ZEROFUNC is semantic.** Prove that the function ZEROFUNC is semantic.

---

**Solution:**
Recall that \( \text{ZEROFUNC}(M) = 1 \) if and only if \( M(x) = 0 \) for every \( x \in \{0,1\}^* \). If \( M \) and \( M' \) are functionally equivalent, then for every \( x, M(x) = M'(x) \). Hence ZEROFUNC\( (M) = 1 \) if and only if \( \text{ZEROFUNC}(M') = 1 \).

---

Often the properties of programs that we are most interested in computing are the semantic ones, since we want to understand the programs’ functionality. Unfortunately, Rice’s Theorem tells us that these properties are all uncomputable:

**Theorem 8.13 — Rice’s Theorem.** Let \( F : \{0,1\}^* \to \{0,1\} \). If \( F \) is semantic and non-trivial then it is uncomputable.

**Proof Idea:**
The idea behind the proof is to show that every semantic non-trivial function \( F \) is at least as hard to compute as HALTONZERO. This will conclude the proof since by Theorem 8.7, HALTONZERO is uncomputable. If a function \( F \) is non trivial then there are two machines \( M_0 \) and \( M_1 \) such that \( F(M_0) = 0 \) and \( F(M_1) = 1 \). So, the goal would be to take a machine \( N \) and find a way to map it into a machine \( M = R(N) \), such that (i) if \( N \) halts on zero then \( M \) is functionally equivalent to \( M_1 \) and (ii) if \( N \) does not halt on zero then \( M \) is functionally equivalent \( M_0 \).

Because \( F \) is semantic, if we achieved this, then we would be guaranteed that HALTONZERO\( (N) = F(R(N)) \), and hence would show that if \( F \) was computable, then HALTONZERO would be computable as well, contradicting Theorem 8.7.

*  

**Proof of Theorem 8.13.** We will not give the proof in full formality, but rather illustrate the proof idea by restricting our attention to a particular semantic function \( F \). However, the same techniques generalize to all possible semantic functions. Define \( \text{MONOTONE} : \{0,1\}^* \to \{0,1\} \) as follows: \( \text{MONOTONE}(M) = 1 \) if there does not exist \( n \in \mathbb{N} \) and two inputs \( x, x' \in \{0,1\}^n \) such that for every \( i \in [n] x_i \leq x'_i \) but \( M(x) \) outputs 1 and \( M(x') = 0 \). That is, \( \text{MONOTONE}(M) = 1 \) if it’s not
possible to find an input \( x \) such that flipping some bits of \( x \) from 0 to 1 will change \( M \)'s output in the other direction from 1 to 0. We will prove that \( \text{MONOTONE} \) is uncomputable, but the proof will easily generalize to any semantic function.

We start by noting that \( \text{MONOTONE} \) is neither the constant zero nor the constant one function:

- The machine \( \text{INF} \) that simply goes into an infinite loop on every input satisfies \( \text{MONOTONE}(\text{INF}) = 1 \), since \( \text{INF} \) is not defined anywhere and so in particular there are no two inputs \( x, x' \) where \( x_i \leq x'_i \) for every \( i \) but \( \text{INF}(x) = 0 \) and \( \text{INF}(x') = 1 \).

- The machine \( \text{PAR} \) that computes the XOR or parity of its input, is not monotone (e.g., \( \text{PAR}(1, 1, 0, 0, \ldots, 0) = 0 \) but \( \text{PAR}(1, 0, 0, \ldots, 0) = 0 \)) and hence \( \text{MONOTONE}(\text{PAR}) = 0 \).

(Note that \( \text{INF} \) and \( \text{PAR} \) are machines and not functions.)

We will now give a reduction from \( \text{HALTONZERO} \) to \( \text{MONOTONE} \). That is, we assume towards a contradiction that there exists an algorithm \( A \) that computes \( \text{MONOTONE} \) and we will build an algorithm \( B \) that computes \( \text{HALTONZERO} \). Our algorithm \( B \) will work as follows:

**Algorithm B:**

**Input:** String \( N \) describing a Turing machine. (Goal: Compute \( \text{HALTONZERO}(N) \))

**Assumption:** Access to Algorithm \( A \) to compute \( \text{MONOTONE} \).

**Operation:**

1. Construct the following machine \( M \): “On input \( z \in \{0, 1\}^* \) do: (a) Run \( M(0) \), (b) Return \( \text{PAR}(z) \)”.
2. Return \( 1 - A(M) \).

To complete the proof we need to show that \( B \) outputs the correct answer, under our assumption that \( A \) computes \( \text{MONOTONE} \). In other words, we need to show that \( \text{HALTONZERO}(N) = 1 - \text{MONOTONE}(M) \). Suppose that \( N \) does not halt on zero. In this case the program \( M \) constructed by Algorithm \( B \) enters into an infinite loop in step (a) and will never reach step (b). Hence in this case \( N \) is functionally equivalent to \( \text{INF} \). (The machine \( N \) is not the same machine as \( \text{INF} \): its description or code is different. But it does have the same input/output behavior (in this case) of never halting on any input. Also, while the program \( M \) will go into an infinite loop on every input, Algorithm \( B \) never actually runs \( M \): it
only produces its code and feeds it to $A$. Hence Algorithm $B$ will 
not enter into an infinite loop even in this case.) Thus in this case, 
$\text{MONOTONE}(N) = \text{MONOTONE}(\text{INF}) = 1$.

If $N$ does halt on zero, then step (a) in $M$ will eventually conclude 
and $M$’s output will be determined by step (b), where it simply out-
puts the parity of its input. Hence in this case, $M$ computes the non-
monotone parity function (i.e., is functionally equivalent to $\text{PAR}$), and 
so we get that $\text{MONOTONE}(M) = \text{MONOTONE}(\text{PAR}) = 0$. In both 
cases, $\text{MONOTONE}(M) = 1 - \text{HALT ON ZERO}(N)$, which is what 
we wanted to prove.

An examination of this proof shows that we did not use anything 
about $\text{MONOTONE}$ beyond the fact that it is semantic and non-trivial. 
For every semantic non-trivial $F$, we can use the same proof, replacing $\text{PAR}$ and $\text{INF}$ with two machines $M_0$ and $M_1$ such that $F(M_0) = 0$ and $F(M_1) = 1$. Such machines must exist if $F$ is non trivial.

\begin{remark}
Semantic is not the same as uncomputable. Rice’s Theorem is so powerful and such a 
popular way of proving uncomputability that people sometimes get confused and think that it is the 
only way to prove uncomputability. In particular, a 
common misconception is that if a function $F$ is not 
semantic then it is computable. This is not at all the 
case.
For example, consider the following function 
$\text{HALT NO YALE} : \{0, 1\}^* \rightarrow \{0, 1\}$. This is a function 
that on input a string that represents a NAND-TM 
program $P$, outputs 1 if and only if both (i) $P$ halts 
on the input 0, and (ii) the program $P$ does not con-
tain a variable with the identifier \text{Yale}. The function 
$\text{HALT NO YALE}$ is clearly not semantic, as it will out-
put two different values when given as input one of 
the following two functionally equivalent programs:

\begin{align*}
\text{Yale}[0] &= \text{NAND}(X[0],X[0]) \\
\text{Y}[0] &= \text{NAND}(X[0],\text{Yale}[0])
\end{align*}

and

\begin{align*}
\text{Harvard}[0] &= \text{NAND}(X[0],X[0]) \\
\text{Y}[0] &= \text{NAND}(X[0],\text{Harvard}[0])
\end{align*}

However, $\text{HALT NO YALE}$ is uncomputable since every 
program $P$ can be transformed into an equivalent 
(and in fact improved :) program $P'$ that does not 
contain the variable \text{Yale}. Hence if we could compute 
$\text{HALT NO YALE}$ then determine halting on zero for 
NAND-TM programs (and hence for Turing machines 
as well).
\end{remark}
Moreover, as we will see in Chapter 10, there are uncomputable functions whose inputs are not programs, and hence for which the adjective “semantic” is not applicable.

Properties such as “the program contains the variable Yale” are sometimes known as syntactic properties. The terms “semantic” and “syntactic” are used beyond the realm of programming languages: a famous example of a syntactically correct but semantically meaningless sentence in English is Chomsky’s “Colorless green ideas sleep furiously.” However, formally defining “syntactic properties” is rather subtle and we will not use this terminology in this book, sticking to the terms “semantic” and “non semantic” only.

8.5.2 Halting and Rice’s Theorem for other Turing-complete models

As we saw before, many natural computational models turn out to be equivalent to one another, in the sense that we can transform a “program” of one model (such as a λ expression, or a game-of-life configurations) into another model (such as a NAND-TM program). This equivalence implies that we can translate the uncomputability of the Halting problem for NAND-TM programs into uncomputability for Halting in other models. For example:

**Theorem 8.15 — NAND-TM Machine Halting.** Let \( \text{NANDTMHALT} : \{0, 1\}^* \rightarrow \{0, 1\} \) be the function that on input strings \( P \in \{0, 1\}^* \) and \( x \in \{0, 1\}^* \) outputs 1 if the NAND-TM program described by \( P \) halts on the input \( x \) and outputs 0 otherwise. Then \( \text{NANDTMHALT} \) is uncomputable.

Once again, this is a good point for you to stop and try to prove the result yourself before reading the proof below.

**Proof.** We have seen in Theorem 6.12 that for every Turing machine \( M \), there is an equivalent NAND-TM program \( P_M \) such that for every \( x, P_M(x) = M(x) \). In particular this means that \( \text{HALT}(M) = \text{NANDTMHALT}(P_M) \).

The transformation \( M \mapsto P_M \) that is obtained from the proof of Theorem 6.12 is constructive. That is, the proof yields a way to compute the map \( M \mapsto P_M \). This means that this proof yields a reduction from the task of computing \( \text{HALT} \) to the task of computing \( \text{NANDTMHALT} \), which means that since \( \text{HALT} \) is uncomputable, neither is \( \text{NANDTMHALT} \).
The same proof carries over to other computational models such as the \( \lambda \) calculus, two dimensional (or even one-dimensional) automata etc. Hence for example, there is no algorithm to decide if a \( \lambda \) expression evaluates the identity function, and no algorithm to decide whether an initial configuration of the game of life will result in eventually coloring the cell \((0,0)\) black or not.

Indeed, we can generalize Rice’s Theorem to all these models. For example, if \( F : \{0,1\}^* \rightarrow \{0,1\} \) is a non-trivial function such that \( F(P) = F(P') \) for every functionally equivalent NAND-TM programs \( P, P' \) then \( F \) is uncomputable, and the same holds for NAND-RAM programs, \( \lambda \)-expressions, and all other Turing complete models (as defined in Definition 7.5), see also Exercise 8.9.

8.5.3 Is software verification doomed? (discussion)

Programs are increasingly being used for mission critical purposes, whether it’s running our banking system, flying planes, or monitoring nuclear reactors. If we can’t even give a certification algorithm that a program correctly computes the parity function, how can we ever be assured that a program does what it is supposed to do? The key insight is that while it is impossible to certify that a general program conforms with a specification, it is possible to write a program in the first place in a way that will make it easier to certify. As a trivial example, if you write a program without loops, then you can certify that it halts. Also, while it might not be possible to certify that an arbitrary program computes the parity function, it is quite possible to write a particular program \( P \) for which we can mathematically prove that \( P \) computes the parity. In fact, writing programs or algorithms and providing proofs for their correctness is what we do all the time in algorithms research.

The field of software verification is concerned with verifying that given programs satisfy certain conditions. These conditions can be that the program computes a certain function, that it never writes into a dangerous memory location, that it respects certain invariants, and others. While the general tasks of verifying this may be uncomputable, researchers have managed to do so for many interesting cases, especially if the program is written in the first place in a formalism or programming language that makes verification easier. That said, verification, especially of large and complex programs, remains a highly challenging task in practice as well, and the number of programs that have been formally proven correct is still quite small. Moreover, even phrasing the right theorem to prove (i.e., the specification) if often a highly non-trivial endeavor.
All functions $F : \{0,1\}^* \to \{0,1\}$

- $F^*$
- $HALT$
- $HALTONZERO$
- $COMPUTES$
- $PARITY$
- ...

Figure 8.8: The set $R$ of computable Boolean functions (Definition 6.4) is a proper subset of the set of all functions mapping $\{0,1\}^*$ to $\{0,1\}$. In this chapter we saw a few examples of elements in the latter set that are not in the former.

### Lecture Recap

- There is a universal Turing machine (or NAND-TM program) $U$ such that on input a description of a Turing machine $M$ and some input $x$, $U(M, x)$ halts and outputs $M(x)$ if (and only if) $M$ halts on input $x$. Unlike in the case of finite computation (i.e., NAND-CIRC programs / circuits), the input to the program $U$ can be a machine $M$ that has more states than $U$ itself.
- Unlike the finite case, there are actually functions that are inherently uncomputable in the sense that they cannot be computed by any Turing machine.
- These include not only some “degenerate” or “esoteric” functions but also functions that people have deeply care about and conjectured that could be computed.
- If the Church-Turing thesis holds then a function $F$ that is uncomputable according to our definition cannot be computed by any means in our physical world.

### 8.6 Exercises

**Exercise 8.1 — NAND-RAM Halt.** Let $NANDRAMHALT : \{0,1\}^* \to \{0,1\}$ be the function such that on input $(P, x)$ where $P$ represents a NAND-RAM program, $NANDRAMHALT(P, x) = 1$ iff $P$ halts on the input $x$. Prove that $NANDRAMHALT$ is uncomputable.

**Exercise 8.2 — Timed halting.** Let $TIMEDHALT : \{0,1\}^* \to \{0,1\}$ be the function that on input (a string representing) a triple $(M, x, T)$,
TimedHALT(M, x, T) = 1 iff the Turing machine M, on input x, halts within at most T steps (where a step is defined as one sequence of reading a symbol from the tape, updating the state, writing a new symbol and (potentially) moving the head).

Prove that TimedHALT is computable.

**Exercise 8.3 — Space halting (challenging).** Let SPACEHALT : \{0, 1\}^* \rightarrow \{0, 1\} be the function that on input (a string representing) a triple \((M, x, T)\), SPACEHALT(M, x, T) = 1 iff the Turing machine M, on input x, halts before its head reached the T-th location of its tape. (We don’t care how many steps M makes, as long as the head stays inside locations \(\{0, \ldots, T-1\}\).

Prove that SPACEHALT is computable. See footnote for hint²

**Exercise 8.4 — Computable compositions.** Suppose that \(F : \{0, 1\}^* \rightarrow \{0, 1\}\) and \(G : \{0, 1\}^* \rightarrow \{0, 1\}\) are computable functions. For each one of the following functions H, either prove that H is necessarily computable or give an example of a pair F and G of computable functions such that H will not be computable. Prove your assertions.

1. \(H(x) = 1\) iff \(F(x) = 1\) OR \(G(x) = 1\).

2. \(H(x) = 1\) iff there exist two nonempty strings \(u, v \in \{0, 1\}^*\) such that \(x = uv\) (i.e., \(x\) is the concatenation of \(u\) and \(v\)), \(F(u) = 1\) and \(G(v) = 1\).

3. \(H(x) = 1\) iff there exist a list \(u_0, \ldots, u_{t-1}\) of non empty strings such that strings \(F(u_i) = 1\) for every \(i \in [t]\) and \(x = u_0u_1\cdots u_{t-1}\).

4. \(H(x) = 1\) iff \(x\) is a valid string representation of a NAND++ program \(P\) such that for every \(z \in \{0, 1\}^*\), on input \(z\) the program \(P\) outputs \(F(z)\).

5. \(H(x) = 1\) iff \(x\) is a valid string representation of a NAND++ program \(P\) such that on input \(x\) the program \(P\) outputs \(F(x)\).

6. \(H(x) = 1\) iff \(x\) is a valid string representation of a NAND++ program \(P\) such that on input \(x\), \(P\) outputs \(F(x)\) after executing at most \(100 \cdot |x|^2\) lines.

**Exercise 8.5** Prove that the following function \(FINITE : \{0, 1\}^* \rightarrow \{0, 1\}\) is uncomputable. On input \(P \in \{0, 1\}^*\), we define \(FINITE(P) = 1\) if and only if \(P\) is a string that represents a NAND++ program such that there only a finite number of inputs \(x \in \{0, 1\}^*\) s.t. \(P(x) = 1\).³

² A machine with alphabet \(\Sigma\) can have at most \(|\Sigma|^T\) choices for the contents of the first \(T\) locations of its tape. What happens if the machine repeats a previously seen configuration, in the sense that the tape contents, the head location, and the current state, are all identical to what they were in some previous state of the execution?

³ Hint: You can use Rice’s Theorem.
Exercise 8.6 — Computing parity. Prove Theorem 8.11 without using Rice’s Theorem.

Exercise 8.7 — TM Equivalence. Let $EQ : \{0, 1\}^* \to \{0, 1\}$ be the function defined as follows, given a string representing a pair $(M, M')$ of Turing machines, $EQ(M, M') = 1$ if $M$ and $M'$ are functionally equivalent as per Definition 8.12. Prove that $EQ$ is uncomputable.

Note that you cannot use Rice’s Theorem directly, as this theorem only deals with functions that take a single Turing machine as input, and $EQ$ takes two machines.

Exercise 8.8 For each of the following two functions, say whether it is computable or not:

1. Given a NAND-TM program $P$, an input $x$, and a number $k$, when we run $P$ on $x$, does the index variable $i$ ever reach $k$?

2. Given a NAND-TM program $P$, an input $x$, and a number $k$, when we run $P$ on $x$, does $P$ ever write to an array at index $k$?

Definition 8.16 — Recursively enumerable. Define a function $F : \{0, 1\}^* \to \{0, 1\}$ to be recursively enumerable if there exists a Turing machine $M$ such that for every $x \in \{0, 1\}^*$, if $F(x) = 1$ then $M(x) = 1$, and if $F(x) = 0$ then $M(x) = \bot$. (i.e., if $F(x) = 0$ then $M$ does not halt on $x$.)

1. Prove that every computable $F$ is also recursively enumerable.

2. Prove that there exists $F$ that is not computable but is recursively enumerable. See footnote for hint. 4

3. Prove that there exists a function $F : \{0, 1\}^* \to \{0, 1\}$ such that $F$ is not recursively enumerable. See footnote for hint. 5

4. Prove that there exists a function $F : \{0, 1\}^* \to \{0, 1\}$ such that $F$ is recursively enumerable but the function $\overline{F}$ defined as $\overline{F}(x) = 1 - F(x)$ is not recursively enumerable. See footnote for hint. 6

::: {exercise title="Rice’s Theorem: standard form" #} In this exercise we will prove Rice’s Theorem in the form that it is typically stated in the literature.

For a Turing machine $M$, define $L(M) \subseteq \{0, 1\}^*$ to be the set of all $x \in \{0, 1\}^*$ such that $M$ halts on the input $x$ and outputs 1. (The set $\text{HALT}$ has this property.

5 You can either use the diagonalization method to prove this directly or show that the set of all recursively enumerable functions is countable.

6 $\text{HALT}$ has this property: show that if both $\text{HALT}$ and $1 - \text{HALT}$ were recursively enumerable then $\text{HALT}$ would be in fact computable.
Show that any $F$ satisfying (b) must be semantic.

Use Theorem 8.13 to prove that for every $F : \{0, 1\}^* \to \{0, 1\}$, if (a) $F$ is neither the constant zero nor the constant one function, and (b) for every $M, M'$ such that $L(M) = L(M')$, $F(M) = F(M')$, then $F$ is uncomputable. See footnote for hint.  

Exercise 8.9 — Rice’s Theorem for general Turing-equivalent models (optional).

Let $\mathcal{F}$ be the set of all partial functions from $\{0, 1\}^*$ to $\{0, 1\}$ and $\mathcal{M} : \{0, 1\}^* \to \mathcal{F}$ be a Turing-equivalent model as defined in Definition 7.5. We define a function $F : \{0, 1\}^* \to \{0, 1\}$ to be $\mathcal{M}$-semantic if there exists some $\mathcal{G} : \mathcal{F} \to \{0, 1\}$ such that $F(P) = \mathcal{G}(\mathcal{M}(P))$ for every $P \in \{0, 1\}^*$.

Prove that for every $\mathcal{M}$-semantic $F : \{0, 1\}^* \to \{0, 1\}$ that is neither the constant one nor the constant zero function, $F$ is uncomputable.

8.7 BIBLIOGRAPHICAL NOTES

The cartoon of the Halting problem in Fig. 8.1 and taken from Charles Cooper’s website.

Section 7.2 in [MM11] gives a highly recommended overview of uncomputability. Gödel, Escher, Bach [Hof99] is a classic popular science book that touches on uncomputability, and unprovability, and specifically Gödel’s Theorem that we will see in Chapter 10. See also the recent book by Holt [Hol18].

The history of the definition of a function is intertwined with the development of mathematics as a field. For many years, a function was identified (as per Euler’s quote above) with the means to calculate the output from the input. In the 1800’s, with the invention of the Fourier series and with the systematic study of continuity and differentiability, people have starting looking at more general kinds of functions, but the modern definition of a function as an arbitrary mapping was not yet universally accepted. For example, in 1899 Poincare wrote “we have seen a mass of bizarre functions which appear to be forced to resemble as little as possible honest functions which serve some purpose. … they are invented on purpose to show that our ancestor’s reasoning was at fault, and we shall never get anything more than that out of them”. Some of this fascinating history is discussed in [grabiner2005the; Gra83; Kle91; Lüt02].

The existence of a universal Turing machine, and the uncomputability of $\text{HALT}$ was first shown by Turing in his seminal paper [Tur37], though closely related results were shown by Church a year before.
These works built on Gödel’s 1931 *incompleteness theorem* that we will discuss in Chapter 10.

Adam Yedidia has written software to help in producing universal Turing machines with a small number of states. This is related to the recreational pastime of “Code Golfing” which is about solving a certain computational task using the as short as possible program.

The diagonalization argument used to prove uncomputability of $F^*$ is derived from Cantor’s argument for the uncountability of the reals discussed in Chapter 2.

Christopher Strachey was an English computer scientist and the inventor of the CPL programming language. He was also an early artificial intelligence visionary, programming a computer to play Checkers and even write love letters in the early 1950’s, see this New Yorker article and this website.

Rice’s Theorem was proven in [Ric53]. It is typically stated in a form somewhat different than what we used, see ??.

We do not discuss in the chapter the concept of recursively enumerable languages, but it is covered briefly in Definition 8.16. As usual, we use function, as opposed to language, notation.
9

Restricted computational models

“Happy families are all alike; every unhappy family is unhappy in its own way”, Leo Tolstoy (opening of the book “Anna Karenina”).

We have seen that many models of computation are Turing equivalent, including Turing machines, NAND-TM/NAND-RAM programs, standard programming languages such as C/Python/Javascript, as well as other models such as the λ calculus and even the game of life. The flip side of this is that for all these models, Rice’s theorem (Theorem 8.13) holds as well, which means that any semantic property of programs in such a model is uncomputable.

The uncomputability of halting and other semantic specification problems for Turing equivalent models motivates restricted computational models that are (a) powerful enough to capture a set of functions useful for certain applications but (b) weak enough that we can still solve semantic specification problems on them. In this chapter we discuss several such examples.

9.1 TURING COMPLETE AND AS A BUG

We have seen that seemingly simple computational models or systems can turn out to be Turing complete. The following webpage lists sev-
eral examples of formalisms that “accidentally” turned out to Turing complete, including supposedly limited languages such as the C preprocessor, CSS, SQL, sendmail configuration, as well as games such as Minecraft, Super Mario, and the card game “Magic: The gathering”. Turing completeness is not always a good thing, as it means that such formalisms can give rise to arbitrarily complex behavior. For example, the postscript format (a precursor of PDF) is a Turing-complete programming language meant to describe documents for printing. The expressive power of postscript can allow for short descriptions of very complex images, but it also gave rise to some nasty surprises, such as the attacks described in this page ranging from using infinite loops as a denial of service attack, to accessing the printer’s file system.

**Example 9.1 — The DAO Hack.** An interesting recent example of the pitfalls of Turing-completeness arose in the context of the cryptocurrency Ethereum. The distinguishing feature of this currency is the ability to design “smart contracts” using an expressive (and in particular Turing-complete) programming language. In our current “human operated” economy, Alice and Bob might sign a contract to agree that if condition $X$ happens then they will jointly invest in Charlie’s company. Ethereum allows Alice and Bob to create a joint venture where Alice and Bob pool their funds together into an account that will be governed by some program $P$ that decides under what conditions it disburses funds from it. For example, one could imagine a piece of code that interacts between Alice, Bob, and some program running on Bob’s car that allows Alice to rent out Bob’s car without any human intervention or overhead.

Specifically Ethereum uses the Turing-complete programming language solidity which has a syntax similar to JavaScript. The flagship of Ethereum was an experiment known as The “Decentralized Autonomous Organization” or The DAO. The idea was to create a smart contract that would create an autonomously run decentralized venture capital fund, without human managers, where shareholders could decide on investment opportunities. The DAO was at the time the biggest crowdfunding success in history. At its height the DAO was worth 150 million dollars, which was more than ten percent of the total Ethereum market. Investing in the DAO (or entering any other “smart contract”) amounts to providing your funds to be run by a computer program. i.e., “code is law”, or to use the words the DAO described itself: “The DAO is borne from immutable, unstoppable, and irrefutable computer code”. Unfortunately, it turns out that (as we saw in Chapter 8) under-
standing the behavior of computer programs is quite a hard thing to do. A hacker (or perhaps, some would say, a savvy investor) was able to fashion an input that caused the DAO code to enter into an infinite recursive loop in which it continuously transferred funds into the hacker’s account, thereby cleaning out about 60 million dollars out of the DAO. While this transaction was “legal” in the sense that it complied with the code of the smart contract, it was obviously not what the humans who wrote this code had in mind. The Ethereum community struggled with the response to this attack. Some tried to the “Robin Hood” approach of using the same loophole to drain the DAO funds into a secure account, but it only had limited success. Eventually, the Ethereum community decided that the code can be mutable, stoppable, and refutable. Specifically, the Ethereum maintainers and miners agreed on a “hard fork” (also known as a “bailout”) to revert history to before the hacker’s transaction occurred. Some community members strongly opposed this decision, and so an alternative currency called Ethereum Classic was created that preserved the original history.

9.2 REGULAR EXPRESSIONS

Searching for a piece of text is a common task in computing. At its heart, the search problem is quite simple. We have a collection \( X = \{x_0, \ldots, x_k\} \) of strings (e.g., files on a hard-drive, or student records in a database), and the user wants to find out the subset of all the \( x \in X \) that are matched by some pattern (e.g., all files whose names end with the string .txt). In full generality, we can allow the user to specify the pattern by specifying a (computable) function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \), where \( F(x) = 1 \) corresponds to the pattern matching \( x \). That is, the user provides a program \( P \) in some Turing-complete programming language such as Python, and the system will return all the \( x \in X \) such that \( P(x) = 1 \). For example, one could search for all text files that contain the string important document or perhaps (letting \( P \) correspond to a neural-network based classifier) all images that contain a cat. However, we don’t want our system to get into an infinite loop just trying to evaluate the program \( P \)!

Because the Halting problem for Turing-complete computational models is uncomputable, we cannot in general verify that a given program \( P \) will halt on a given input. For this reason, typical systems for searching files or databases do not allow users to specify the patterns using full-fledged programming languages. Rather, such systems use restricted computational models that on the one hand are rich enough to capture many of the queries needed in practice (e.g., all filenames
ending with .txt, or all phone numbers of the form (617) xxx-xxxx), but on the other hand are restricted enough so that they cannot result in an infinite loop.

One of the most popular such computational models is regular expressions. If you ever used an advanced text editor, a command line shell, or have done any kind of manipulations of text files, then you have probably come across regular expressions.

A regular expression over some alphabet $\Sigma$ is obtained by combining elements of $\Sigma$ with the operation of concatenation, as well as $|$ (corresponding to or) and $*$ (corresponding to repetition zero or more times). (Common implementations of regular expressions in programming languages and shells typically include some extra operations on top of $|$ and $*$, but these operations can be implemented as “syntactic sugar” using the operators $|$ and $*$.) For example, the following regular expression over the alphabet $\{0, 1\}$ corresponds to the set of all strings $x \in \{0, 1\}^*$ where every digit is repeated at least twice:

$$((00^* 11^*))^*.$$ (9.1)

The following regular expression over the alphabet $\{a, \ldots, z, 0, \ldots, 9\}$ corresponds to the set of all strings that consist of a sequence of one or more of the letters $a$-$d$ followed by a sequence of one or more digits (without a leading zero):

$$(a|b|c|d)(a|b|c|d)^*(1|2|3|4|5|6|7|8|9)(0|1|2|3|4|5|6|7|8|9)^*.$$ (9.2)

Formally, regular expressions are defined by the following recursive definition:

**Definition 9.2 — Regular expression.** A regular expression $e$ over an alphabet $\Sigma$ is a string over $\Sigma$ $\cup \{(), |, *, \emptyset, ""\}$ that has one of the following forms:

1. $e = \sigma$ where $\sigma \in \Sigma$
2. $e = (e'|e")$ where $e’, e"$ are regular expressions.
3. $e = (e')(e")$ where $e’, e"$ are regular expressions. (We often drop the parenthesis when there is no danger of confusion and so write this as $e’ e\"$.)
4. $e = (e')^*$ where $e’$ is a regular expression.
Finally we also allow the following “edge cases”: \( e = \emptyset \) and \( e = "" \). These are the regular expressions corresponding to accepting no strings, and accepting only the empty string respectively.

We will drop parenthesis when they can be inferred from the context. We also use the convention that OR and concatenation are left-associative, and give higher precedence to \( * \), then concatenation, and then OR. Thus for example we write \( 00^*|11 \) instead of \((((0)(0^*))|((1)(1)))\).

Every regular expression \( e \) corresponds to a function \( \Phi_e : \Sigma^* \to \{0, 1\} \) where \( \Phi_e(x) = 1 \) if \( x \) matches the regular expression. For example, if \( e = (00|11)^* \) then \( \Phi_e(110011) = 1 \) but \( \Phi_e(101) = 0 \) (can you see why?).

The formal definition of \( \Phi_e \) is one of those definitions that is more cumbersome to write than to grasp. Thus it might be easier for you to first work it out on your own and then check that your definition matches what is written below.

**Definition 9.3 — Matching a regular expression.** Let \( e \) be a regular expression over the alphabet \( \Sigma \). The function \( \Phi_e : \Sigma^* \to \{0, 1\} \) is defined as follows:

1. If \( e = \sigma \) then \( \Phi_e(x) = 1 \) iff \( x = \sigma \).
2. If \( e = (e' | e'') \) then \( \Phi_e(x) = \Phi_{e'}(x) \lor \Phi_{e''}(x) \) where \( \lor \) is the OR operator.
3. If \( e = (e')(e'') \) then \( \Phi_e(x) = 1 \) iff there is some \( x', x'' \in \Sigma^* \) such that \( x \) is the concatenation of \( x' \) and \( x'' \) and \( \Phi_{e'}(x') = \Phi_{e''}(x'') = 1 \).
4. If \( e = (e')^* \) then \( \Phi_e(x) = 1 \) iff there are is \( k \in \mathbb{N} \) and some \( x_0, \ldots, x_{k-1} \in \Sigma^* \) such that \( x \) is the concatenation \( x_0 \cdots x_{k-1} \) and \( \Phi_{e'}(x_i) = 1 \) for every \( i \in [k] \).
5. Finally, for the edge cases \( \Phi_\emptyset \) is the constant zero function, and \( \Phi_{""} \) is the function that only outputs 1 on the empty string "".

We say that a regular expression \( e \) over \( \Sigma \) matches a string \( x \in \Sigma^* \) if \( \Phi_e(x) = 1 \). We say that a function \( F : \Sigma^* \to \{0, 1\} \) is regular if \( F = \Phi_e \) for some regular expression \( e \).\(^1\)

\(^1\) We use function notation in this book, but other texts often use the notion of languages, which are sets of strings. In that notation a language \( L \subseteq \Sigma^* \) is called regular if and only if the corresponding function \( F_L \) is regular, where \( F_L : \Sigma^* \to \{0, 1\} \) is the function that outputs 1 on \( x \) iff \( x \in L \).
The definitions above are not inherently difficult, but are a bit cumbersome. So you should pause here and go over it again until you understand why it corresponds to our intuitive notion of regular expressions. This is important not just for understanding regular expressions themselves (which are used time and again in a great many applications) but also for getting better at understanding recursive definitions in general.

Example 9.4 — A regular function. Let $\Sigma = \{a, b, c, d, 0, 1, 2, 3, 4, 5, 6, 7, 8, 9\}$ and $F : \Sigma^* \rightarrow \{0, 1\}$ be the function such that $F(x)$ outputs 1 iff $x$ consists of one or more of the letters $a$-$d$ followed by a sequence of one or more digits (without a leading zero). Then $F$ is a regular function, since $F = \Phi_e$ where

$$
e = (a|b|c|d)(a|b|c|d)^*(0|1|2|3|4|5|6|7|8|9)(0|1|2|3|4|5|6|7|8|9)^* \quad (9.3)$$

is the expression we saw in (9.2).

If we wanted to verify, for example, that $\Phi_e(abc12078) = 1$, we can do so by noticing that the expression $(a|b|c|d)$ matches the string $a$, $(a|b|c|d)^*$ matches $bc$, $(0|1|2|3|4|5|6|7|8|9)$ matches the string $1$, and the expression $(0|1|2|3|4|5|6|7|8|9)^*$ matches the string $2078$. Each one of those boils down to a simpler expression. For example, the expression $(a|b|c|d)^*$ matches the string $bc$ because both of the one-character strings $b$ and $c$ are matched by the expression $a|b|c|d$.

Regular expression can be defined over any finite alphabet $\Sigma$, but as usual, we will focus our attention on the binary case, where $\Sigma = \{0, 1\}$. Most (if not all) of the theoretical and practical general insights about regular expressions can be gleaned from studying the binary case.

We can think of regular expressions as a type of “programming language”. That is, we can think of a regular expression $e$ over the alphabet $\Sigma$ as a program that computes the function $\Phi_e : \Sigma^* \rightarrow \{0, 1\}$. (You can also think of regular expressions as generative models, since you can think of them as giving a recipe how to generate strings that match them.) This “regular expression programming language” is simpler than general programming languages, in the sense that for every regular expression $e$, the function $\Phi_e$ is computable (and so in particular can be evaluated by an always-halting Turing machine).
Theorem 9.5 — Regular expression always halt. For every regular expression \( e \) over \( \{0, 1\} \), the function \( \Phi_e : \{0, 1\}^* \rightarrow \{0, 1\} \) is computable.
That is, there is a Turing machine $M$ such that for every $x \in \{0, 1\}^*$, on input $x$, $M$ halts with the output $\Phi_e(x)$.

We state Theorem 9.5 for regular expressions over the binary alphabet $\{0, 1\}$, but it generalizes to any finite alphabet $\Sigma$.

**Proof Idea:**

The proof relies on the observation that Definition 9.3 actually specifies a recursive algorithm for computing $\Phi_e$. Specifically, each one of our operations - concatenation, OR, and star - can be thought of as reducing the task of testing whether an expression $e$ matches a string $x$ to testing whether some sub-expressions of $e$ match substrings of $x$. Since these sub-expressions are always shorter than the original expression, this yields a recursive algorithm for checking if $e$ matches $x$ which will eventually terminate at the base cases of the expressions that correspond to a single symbol or the empty string.

Proof of Theorem 9.5. Definition 9.3 gives a way of recursively computing $\Phi_e$. The key observation is that in our recursive definition of regular expressions, whenever $e$ is made up of one or two expressions $e', e''$ then these two regular expressions are smaller than $e$, and eventually (when they have size 1) then they must correspond to the non-recursive case of a single alphabet symbol.

Therefore, we can prove the theorem by induction over the length $m$ of $e$ (i.e., the number of symbols in the string $e$, also denoted as $|e|$). For $m = 1$, $e$ is either a single alphabet symbol, "" or $\emptyset$, and so computing the function $\Phi_e$ is straightforward. In the general case, for $m = |e|$ we assume by the induction hypothesis that we have proven the theorem for all expressions of length smaller than $m$. Now, such an expression of length larger than one can obtained one of three cases using the OR, concatenation, or star operations. We now show that $\Phi_e$ will be computable in all these cases:

**Case 1:** $e = (e'|e'')$ where $e', e''$ are shorter regular expressions.

In this case by the inductive hypothesis we can compute $\Phi_{e'}$ and $\Phi_{e''}$ and so can compute $\Phi_e(x)$ as $\Phi_{e'}(x) \lor \Phi_{e''}(x)$ (where $\lor$ is the OR operator).

**Case 2:** $e = (e')(e'')$ where $e', e''$ are regular expressions.

In this case by the inductive hypothesis we can compute $\Phi_{e'}$ and $\Phi_{e''}$ and so can compute $\Phi_e(x)$ as

$$
\bigvee_{i=0}^{|x|-1} (\Phi_{e'}(x_0 \cdots x_{i-1}) \land \Phi_{e''}(x_i \cdots x_{|x|-1}))
$$

(9.4)

where $\land$ is the AND operator and for $i < j$, $x_j \cdots x_i$ refers to the empty string.
Case 3: $e = (e')^*$ where $e'$ is a regular expression.

In this case by the inductive hypothesis we can compute $\Phi_{e'}$ and so we can compute $\Phi_e(x)$ by enumerating over all $k$ from 1 to $|x|$, and all ways to write $x$ as the concatenation of $k$ nonempty strings $x_0 \cdots x_{k-1}$ (we can do so by enumerating over all possible $k-1$ positions in which one string stops and the other begins). If for one of those partitions, $\Phi_{e'}(x_0) = \cdots = \Phi_{e'}(x_{k-1}) = 1$ then we output 1. Otherwise we output 0. We can restrict attention to partitions of $x$ as $x = x_0 \cdots x_{k-1}$ where all the $x_i$’s are nonempty since if some of the $x_i$’s are empty we can simply drop them and still be left with a valid partition.

These three cases exhaust all the possibilities for an expression of length larger than one, and hence this completes the proof.

9.3 DETERMINISTIC FINITE AUTOMATA, AND EFFICIENT MATCHING OF REGULAR EXPRESSIONS (OPTIONAL)

The proof of Theorem 9.5 gives a recursive algorithm to evaluate whether a given string matches or not a regular expression. But it is not a very efficient algorithm.

However, it turns out that there is a much more efficient algorithm that can match regular expressions in linear (i.e., $O(n)$) time. Since we have not yet covered the topics of time and space complexity, we describe this algorithm in high level terms, without making the computational model precise, using the colloquial notion of $O(n)$ running time as is used in introduction to programming courses and whiteboard coding interviews. We will see a formal definition of time complexity in Chapter 12.

**Theorem 9.6 — Matching regular expressions in linear time.** Let $e$ be a regular expression. Then there is an $O(n)$ time algorithm that computes $\Phi_e$.

The implicit constant in the $O(n)$ term of Theorem 9.6 depends on the expression $e$. Thus, another way to state Theorem 9.6 is that for every expression $e$, there is some constant $c$ and an algorithm $A$ that computes $\Phi_e$ on $n$-bit inputs using at most $c \cdot n$ steps. This makes sense, since in practice we often want to compute $\Phi_e(x)$ for a small regular expression $e$ and a large document $x$. Theorem 9.6 tells us that we can do so with running time that scales linearly with the size of the document, even if it has (potentially) worse dependence on the size of the regular expression.

**Proof Idea:**

The idea is to define a more efficient recursive algorithm, that determines whether $e$ matches a string $x \in \{0, 1\}^n$ by reducing this task...
to determining whether a related expression $e'$ matches $x_0, \ldots, x_{n-1}$.
This will result in an expression for the running time of the form
$T(n) = T(n - 1) + O(1)$ which solves to $T(n) = O(n)$.

Proof of Theorem 9.6. The central definition for this proof is the notion of a restriction of a regular expression. The idea is that for every regular expression $e$ and symbol $\sigma$ in its alphabet, it is possible to define a regular expression $e[\sigma]$ such that $e[\sigma]$ matches a string $x$ if and only if $e$ matches the string $x\sigma$. For example, if $e$ is the regular expression $01|(01)^*(01)$ (i.e., one or more occurrences of $01$), then $e[1]$ is equal to $(0)(01)^*0$ and $e[0]$ will be $\emptyset$. (Can you see why?)

For simplicity, from now on we fix our attention to the case that the alphabet $\Sigma$ is $\{0, 1\}$. Given a regular expression $e$ and $\sigma \in \{0, 1\}$, we can compute $e[\sigma]$ recursively as follows:

1. If $e$ consists of a single symbol (i.e. $e = \tau$ for $\tau \in \{0, 1\}$) then $e[\sigma] = **$ if $\tau = \sigma$ and $e[\sigma] = \emptyset$ otherwise.
2. If $e = e'|e''$ then $e[\sigma] = e'[\sigma]|e''[\sigma]$.
3. If $e = e'|e''$ then $e[\sigma] = e'|e''[\sigma]$ if $e''$ can not match the empty string.
   Otherwise, $e[\sigma] = e'|e''[\sigma] \mid e'[\sigma]$
4. If $e = (e')^*$ then $e[\sigma] = (e')^*(e'[\sigma])$.
5. If $e = **$ or $e = \emptyset$ then $e[\sigma] = \emptyset$.

By checking all these cases, one can verify that it is indeed the case that for every regular expression $e$, $\sigma \in \{0, 1\}$ and $x \in \{0, 1\}^*$, $e[\sigma]$ matches $x$ if and only if $e$ matches $x\sigma$. We let $C(\ell)$ denote the time to compute $e[\sigma]$ for regular expressions of length at most $\ell$. The value $C(\ell)$ can be shown to be polynomial in $\ell$, though this is not important for this theorem, since we only care about the dependence of the time to compute $\Phi_e(x)$ on the length of $x$ and not about the dependence of this time on the length of $e$.

Using this notion of restriction, we can define the following recursive algorithm for regular expression matching:
Algorithm 9.7 — Regular expression matching in linear time.

**Input:** Regular expression $e$ over $\{0, 1\}$ and $x \in \{0, 1\}^n$ for $n \in \mathbb{N}$.

**Goal:** Compute $\Phi_e(x)$

**Operation:**

1. If $x = \"\"$ then return 1 if and only if $\Phi_e(\"\") = 1$. (This can be either computed directly or using the algorithm of Theorem 9.5 in time which is a constant depending only on the regular expression $e$.)
2. Otherwise, compute $\Phi_{e[x_{n-1}]}(x_0 \cdots x_{n-1})$ recursively and output the result.

By the definition of a restriction, for every $\sigma \in \{0, 1\}$ and $x' \in \{0, 1\}^*$, the expression $e$ matches $x'\sigma$ if and only if $e[\sigma]$ matches $x'$. Hence for every $e$ and $x \in \{0, 1\}^n$, $\Phi_{e[x_{n-1}]}(x_0 \cdots x_{n-2}) = \Phi_e(x)$ and Algorithm 9.7 does return the correct answer. The only remaining task is to analyze its running time.

Algorithm 9.7 is a recursive algorithm that on input an expression $e$ and a string $x \in \{0, 1\}^n$, does some constant time computation and then calls itself on input some expression $e'$ and a string $x$ of length $n - 1$. It will terminate after $n$ steps when it reaches a string of length 0. So, to calculate the running time of Algorithm 9.7 we need to analyze the cost of each step.

Specifically, the running time $T(e, n)$ that it takes for Algorithm 9.7 to compute $\Phi_e$ for inputs of length $n$ satisfies the recursive equation:

$$T(e, n) = \max\{T(e[0], n - 1), T(e[1], n - 1)\} + C(|e|) \quad (9.5)$$

where $C(\ell)$, as before, denotes the time to compute $e[\sigma]$ for expressions $e$ of length at most $\ell$. (In the base case $n = 0$, $T(e, 0)$ is equal to some constant depending only on $e$.)

To get some intuition for the expression Eq. (9.5), let us open up the recursion for one level, writing $T(e, n)$ as

$$T(e, n) = \max\{T(e[0][0], n - 2) + C(|e[0]|),\ T(e[0][1], n - 2) + C(|e[0]|),\ T(e[1][0], n - 2) + C(|e[1]|),\ T(e[1][1], n - 2) + C(|e[1]|)\} + C(|e|). \quad (9.6)$$

Continuing this way, we can see that $T(e, n) \leq n \cdot C(\ell) + O(1)$ where $\ell$ is the largest length of any expression $e'$ that we encounter along the way. Therefore, the following claim suffices to show that Algorithm 9.7 runs in linear time.
Claim: Let $e$ be a regular expression over $\{0, 1\}$, then there is some constant $c$ such that for every string $\alpha \in \{0, 1\}^*$, if we restrict $e$ to $\alpha_0$, and then to $\alpha_1$ and so on and so forth, the resulting expression has length at most $c$.

Proof of claim: For a regular expression $e$ over $\{0, 1\}$ and $\alpha \in \{0, 1\}^m$, we denote by $e[\alpha]$ the expression $e[\alpha_0][\alpha_1] \cdots [\alpha_{m-1}]$ obtained by restricting $e$ to $\alpha_0$ and then to $\alpha_1$ and so on. We let $S(e) = \{e[\alpha]|\alpha \in \{0, 1\}^*\}$. We will prove the claim by showing that for every $e$, the set $S(e)$ is finite, and hence so is the number $c(e)$ which is the maximum length of $e'$ for $e' \in S(e)$.

We prove this by induction on the structure of $e$. If $e$ is a symbol, the empty string, or the empty set, then this is straightforward to show as the most expressions $S(e)$ can contain are the expression itself, "", and $\emptyset$. Otherwise we split to the two cases (i) $e = e' \ast$ and (ii) $e = e'' e''$, where $e'$, $e''$ are smaller expressions (and hence by the induction hypothesis $S(e')$ and $S(e'')$ are finite). In the case (i), if $e = (e')\ast$ then $e[\alpha]$ is either equal to $(e')\ast e'[\alpha]$ or it is simply the empty set if $e'[\alpha] = \emptyset$. Since $e'[\alpha]$ is in the set $S(e')$, the number of distinct expressions in $S(e)$ is at most $|S(e')| + 1$. In the case (ii), if $e = e'' e''$ then all the restrictions of $e$ to strings $\alpha$ will either have the form $e''[\alpha]$ or the form $e''[\alpha] e'[\alpha']$ where $\alpha'$ is some string such that $\alpha = \alpha' \alpha''$ and $e[\alpha']$ matches the empty string. Since $e''[\alpha] \in S(e'')$ and $e'[\alpha'] \in S(e')$, the number of the possible distinct expressions of the form $e[\alpha]$ is at most $|S(e'')| + |S(e'')| \cdot |S(e')|$. This completes the proof of the claim.

The bottom line is that while running Algorithm 9.7 on a regular expression $e$, all the expressions we ever encounter are in the finite set $S(e)$, no matter how large the input $x$ is, and so the running time of Algorithm 9.7 satisfies the equation $T(n) = T(n - 1) + C'$ for some constant $C'$ depending on $e$. This solves to $O(n)$ where the implicit constant in the Oh notation can (and will) depend on $e$ but crucially, not on the length of the input $x$.

9.3.1 Matching regular expressions using constant memory

Theorem 9.6 is already quite impressive, but we can do even better. Specifically, no matter how long the string $x$ is, we can compute $\Phi_e(x)$ by maintaining only a constant amount of memory and moreover
making a single pass over $x$. That is, the algorithm will scan the input $x$ once from start to finish, and then determine whether or not $x$ is matched by the expression $e$. This is important in the common case of trying to match a short regular expression over a huge file or document that might not even fit in our computer’s memory. A single-pass constant-memory algorithm is also known as a deterministic finite automaton (DFA) (see Section 9.3.2). There is a beautiful theory on the properties of DFA’s and their connections with regular expressions. In particular, as we’ll see in Theorem 9.12, a function is regular if and only if it can be computed by a DFA. We start with showing the “only if” direction:

**Theorem 9.8 — DFA for regular expression matching.** Let $e$ be a regular expression. Then there is an algorithm that on input $x \in \{0, 1\}^*$ computes $\Phi_e(x)$ while making a single pass over $x$ and maintaining a constant amount of memory.

**Proof Idea:**

The idea is to replace the recursive algorithm of Algorithm 9.7 with a dynamic program, using the technique of memoization. If you haven’t taken yet an algorithms course, you might not know these techniques. This is OK; while this more efficient algorithm is crucial for the many practical applications of regular expressions, it is not of great importance for this book.

*Proof of Theorem 9.8.* We will replace the recursive Algorithm 9.7 with the following iterative algorithm:
Algorithm 9.9 — Constant memory regular expression matching.

**Input:** Regular expression $e$ over $\{0, 1\}$, string $x \in \{0, 1\}^n$.

**Goals:** Compute $\Phi_e(x)$.

**Operation:**

1. Let $S = S(e)$ be the set $\{e[\alpha] | \alpha \in \{0, 1\}^*\}$ as defined in the proof of Theorem 9.6. Note that $S$ is finite and by definition, for every $e' \in S$ and $\sigma \in \{0, 1\}$, $e'[\sigma]$ is in $S$ as well.

2. Define a Boolean variable $v_{e'}$ for every $e' \in S$. Initially we set $v_{e'} = 1$ if and only if $e'$ matches the empty string.

3. For $i = 0, \ldots, n - 1$ do the following:
   
   a. *Copy the variables $\{v_{e'}\}$ to temporary variables:* For every $e' \in S$, we set $\text{temp}_{e'} = v_{e'}$.
   
   b. *Update the variables $\{v_{e'}\}$ based on the $i$-th bit of $x$: Let $\sigma = x_i$ and set $v_{e'} = \text{temp}_{e'[\sigma]}$ for every $e' \in S$.

4. Output $v_e$.

Algorithm 9.9 maintains the invariant that at the end of step $i$, for every $e' \in S$, the variable $v_{e'}$ is equal if and only if $e'$ matches the string $x_0 \cdots x_{i-1}$. In particular, at the very end, $v_e$ is equal to 1 if and only if $e$ matches the full string $x_0 \cdots x_{n-1}$. Algorithm 9.9 only maintains a constant number of variables (as $S$ is finite), and that it proceeds in one linear scan over the input, and so this proves the theorem.

9.3.2 Deterministic Finite Automata

In Computer Science, a single-pass constant-memory algorithm is also known as a *Deterministic Finite Automaton* or DFA (another name for DFA's is a *finite state machine*). That is, we can think of such an algorithm as a “machine” that can be in one of $C$ states, for some constant $C$. The machine starts in some initial state, and then reads its input $x \in \{0, 1\}^*$ one bit at a time. Whenever the machine reads a bit $\sigma \in \{0, 1\}$, it transitions into a new state based on $\sigma$ and its prior state. The output of the machine is based the final state. Every constant-memory one-pass algorithm corresponds to such a machine. If an algorithm uses $c$ bits of memory, then the contents of its memory are a string of length $c$. Since there are $2^c$ such strings, at any point in the execution, such an algorithm can be in one of $2^c$ states.
Example 9.10 — DFA for XOR. Here is a DFA for computing the function $XOR : \{0, 1\}^n \rightarrow \{0, 1\}$ that maps $x$ to $\sum_{i \in [n]} x_i \mod 2$.

We will have two states: 0 and 1. The set of accepting states is $\{1\}$, and if we are in a state $v \in \{0, 1\}$ and read the bit $\sigma$, we will transition to the state $v$ if $\sigma = 0$ and to the state $1 - v$ if $\sigma = 1$. In other words, we transition to the state $v \oplus \sigma$. Hence we can think of this algorithm's execution on input $x \in \{0, 1\}^n$ as follows:

- Let $v_t$ be the state of the automaton at step $t$. We initialize $v_0 = 0$.
- For every $i \in [n]$, let $v_{i+1} = v_i \oplus x_i$.
- Output $v_n$.

You can verify that the output of this algorithm is $x_0 \oplus x_1 \oplus \cdots \oplus x_{n-1} = XOR(x)$. We can also describe this DFA graphically, see Fig. 9.2.

The formal definition of a DFA is the following:

Definition 9.11 — Deterministic Finite Automaton. A deterministic finite automaton (DFA) with $C$ states over $\{0, 1\}$ is a pair $(T, S)$ with $T : [C] \times \{0, 1\} \rightarrow [C]$ and $S \subseteq [C]$. The function $T$ is known as the transition function of the DFA and the set $S$ is known as the set of accepting states.

We say that $(T, S)$ computes a function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ if for every $n \in \mathbb{N}$ and $x \in \{0, 1\}^n$, if we define $v_0 = 0$ and $v_{i+1} = T(v_i, x_i)$ for every $i \in [n]$, then

$$v_n \in S \iff F(x) = 1$$ (9.7)

Our treatment of automata in this book is quite brief. If you find this definition confusing, there are plenty of resources that help you get more comfortable with DFA’s. In particular, Chapter 1 of Sipser’s book [Sip97] contains an excellent exposition of this material. There are also many websites with online simulators for automata, as well as translators from regular expressions to automata and vice versa (see for example here and here).

Sipser defines a DFAs as a five-tuple $(Q, \Sigma, \delta, q_0, F)$ where $Q$ is the set of states, $\Sigma$ is the alphabet, $\delta$ is the transition function, $q_0$ is the initial state, and $F$ is the set of accepting states. In this book the set of states is always of the form $Q = \{0, \ldots, C - 1\}$ and the initial
The following theorem is the central result of automata theory:

**Theorem 9.12 — DFA and regular expression equivalency.** Let $F : \{0, 1\}^* \to \{0, 1\}$. Then $F$ is regular if and only if there exists a DFA $(T, S)$ that computes $F$.

**Proof Idea:**

One direction follows from Theorem 9.8, which shows that for every regular expression $e$, the function $\Phi_e$ can be computed by a DFA (see for example Fig. 9.3). For the other direction, we show that given a DFA $(T, S)$ for every $v, w \in [C]$ we can find a regular expression that would match $x \in \{0, 1\}^*$ if an only if the DFA starting in state $v$, will end up in state $w$ after reading $x$.

**Proof of Theorem 9.12.** Since Theorem 9.8 proves the “only if” direction, we only need to show the “if” direction. Let $A = (T, S)$ be a DFA with $C$ states that computes the function $F$. We need to show that $F$ is regular.

For every $v, w \in [C]$, we let $F_{v,w} : \{0, 1\}^* \to \{0, 1\}$ be the function that maps $x \in \{0, 1\}^*$ to 1 if and only if the DFA $A$, starting at the state $v$, will reach the state $w$ if it reads the input $x$. We will prove that $F_{v,w}$ is regular for every $v, w$. This will prove the theorem, since by Definition 9.11, $F(x)$ is equal to the OR of $F_{0,w}(x)$ for every $w \in S$. Hence if we have a regular expression for every function of the form $F_{v,w}$ then (using the | operation) we can obtain a regular expression for $F$ as well.

To give regular expressions for the functions $F_{v,w}$ we start by defining the following functions $F_{v,w}^t$: for every $v, w \in [C]$ and $0 \leq t \leq C$, $F_{v,w}^t(x) = 1$ if and only if starting from $v$ and observing $x$, the automata reaches $w$ with all intermediate states in the set $[t] = \{0, \ldots, t - 1\}$ (see Fig. 9.4). That is, while $v, w$ themselves might be outside $[t]$, $F_{v,w}^t(x) = 1$ if and only if throughout the execution of the automaton on the input $x$ (when initiated at $v$) it never enters any of the states outside $[t]$ and still ends up at $w$. If $t = 0$ then $[t]$ is the empty set, and hence $F_{v,w}^0(x) = 1$ if and only if the automaton reaches $w$ from $v$ directly on $x$, without any intermediate state. If $t = C$ then all states are in $[t]$, and hence $F_{v,w}^C = F_{v,w}$.
We will prove the theorem by induction on \( t \), showing that \( F^t_{v,w} \) is regular for every \( v, w \) and \( t \). For the base case of \( t = 0 \), \( F^0_{v,w} \) is regular for every \( v, w \) since it can be described one of the expressions "", \( \emptyset \), 0, 1 or \( 0|1 \). Specifically, if \( v = w \) then \( F^0_{v,w}(x) = 1 \) if and only if \( x \) is the empty string. If \( v \neq w \) then \( F^0_{v,w}(x) = 1 \) if and only if \( x \) consists of a single symbol \( \sigma \in \{0, 1\} \) and \( T(v, \sigma) = w \). Therefore in this case \( F^0_{v,w} \) corresponds to one of the four regular expressions \( 0|1 \), \( 0 \), \( 1 \) or \( \emptyset \), depending on whether \( A \) transitions to \( w \) from \( v \) when it reads either 0 or 1, only one of these symbols, or neither.

**Inductive step:** Now that we’ve seen the base case, let’s prove the general case by induction. Assume, via the induction hypothesis, that for every \( v', w' \in [C] \), we have a regular expression \( R^t_{v',w'} \) that computes \( F^t_{v',w'} \). We need to prove that \( F^{t+1}_{v,w} \) is regular for every \( v, w \). If the automaton arrives from \( v \) to \( w \) using the intermediate states \([t + 1]\), then it visits the \( t \)-th state zero or more times. If the path labeled by \( x \) causes the automaton to get from \( v \) to \( w \) without visiting the \( t \)-th state at all, then \( x \) is matched by the regular expression \( R^t_{v,w} \). If the path labeled by \( x \) causes the automaton to get from \( v \) to \( w \) while visiting the \( t \)-th state \( k > 0 \) times then we can think of this path as:

- First travel from \( v \) to \( t \) using only intermediate states in \([t - 1]\).
- Then go from \( t \) back to itself \( k - 1 \) using only intermediate states in \([t - 1]\).
- Then go from \( t \) to \( w \) using only intermediate states in \([t - 1]\).

Therefore in this case the string \( x \) is matched by the regular expression \( R^t_{v,t}(R^t_{t,t})^* R^t_{t,w} \). (See also Fig. 9.5.) Therefore we can compute \( F^{t+1}_{v,w} \) using the regular expression

\[
R^t_{v,w} \mid R^t_{v,t}(R^t_{t,t})^* R^t_{t,w}.
\] (9.8)

This completes the proof of the inductive step and hence of the theorem.

9.3.3 **Regular functions are closed under complement**

Here is an important corollary of Theorem 9.12:

**Lemma 9.13 — Regular expressions closed under complement.** If \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) is regular then so is the function \( \overline{F} \), where \( \overline{F}(x) = 1 - F(x) \) for every \( x \in \{0, 1\}^* \).

**Proof.** If \( F \) is regular then by Theorem 9.6 it can be computed by a constant-space one-pass algorithm \( A \). But then the algorithm \( \overline{A} \) which does the same computation and outputs the negation of the output
of $A$ also utilizes constant space and one pass and computes $\overline{F}$. By Theorem 9.12 this implies that $\overline{F}$ is regular as well.

### 9.4 LIMITATIONS OF REGULAR EXPRESSIONS

The fact that functions computed by regular expressions always halt is one of the reasons why they are so useful. When you make a regular expression search, you are guaranteed that it will terminate with a result. This is why operating systems and text editors often restrict their search interface to regular expressions and don’t allow searching by specifying an arbitrary function. But this always-halting property comes at a cost. Regular expressions cannot compute every function that is computable by Turing machines. In fact there are some very simple (and useful!) functions that they cannot compute. Here is one example:

**Lemma 9.14 — Matching parenthesis.** Let $\Sigma = \{\langle, \rangle\}$ and $\text{MATCHPAREN} : \Sigma^* \rightarrow \{0, 1\}$ be the function that given a string of parenthesis, outputs 1 if and only if every opening parenthesis is matched by a corresponding closed one. Then there is no regular expression over $\Sigma$ that computes $\text{MATCHPAREN}$.

**Lemma 9.14** is a consequence of the following result, which is known as the **pumping lemma**:

**Theorem 9.15 — Pumping Lemma.** Let $e$ be a regular expression over some alphabet $\Sigma$. Then there is some number $n_0$ such that for ev-
Every $w \in \{0, 1\}^*$ with $|w| > n_0$ and $\Phi_e(w) = 1$, we can write $w = xyz$ for strings $x, y, z \in \Sigma^*$ satisfying the following conditions:

1. $|y| \geq 1$.
2. $|xy| \leq n_0$.
3. $\Phi_e(xyz^k) = 1$ for every $k \in \mathbb{N}$.

**Proof Idea:**

The idea behind the proof is the following. Let $n_0$ be twice the number of symbols that are used in the expression $e$, then the only way that there is some $w$ with $|w| > n_0$ and $\Phi_e(w) = 1$ is that $e$ contains the $*$ (i.e. star) operator and that there is a nonempty substring $y$ of $w$ that was matched by $(e')^*$ for some sub-expression $e'$ of $e$. We can now repeat $y$ any number of times and still get a matching string. See also Fig. 9.6.

**Proof of Theorem 9.15.** To prove the lemma formally, we use induction on the length of the expression. Like all induction proofs, this is going to be somewhat lengthy, but at the end of the day it directly follows.
the intuition above that somewhere we must have used the star operation. Reading this proof, and in particular understanding how the formal proof below corresponds to the intuitive idea above, is a very good way to get more comfortable with inductive proofs of this form.

Our inductive hypothesis is that for an \( n \) length expression, \( n_0 = 2n \) satisfies the conditions of the lemma. The base case is when the expression is a single symbol \( \sigma \in \Sigma \) or that the expression is \( \emptyset \) or "". In all these cases the conditions of the lemma are satisfied simply because there \( n_0 = 2 \) and there is no string \( x \) of length larger than \( n_0 \) that is matched by the expression.

We now prove the inductive step. Let \( e \) be a regular expression with \( n > 1 \) symbols. We set \( n_0 = 2n \) and let \( w \in \Sigma^* \) be a string satisfying \( |w| > n_0 \). Since \( e \) has more than one symbol, it has one of the the forms (a) \( e'|e'' \), (b) \( e'(e'')(e'') \), or (c) \( (e')^* \) where in all these cases the subexpressions \( e' \) and \( e'' \) have fewer symbols than \( e \) and hence satisfy the induction hypothesis.

In the case (a), every string \( w \) matched by \( e \) must be matched by either \( e' \) or \( e'' \). If \( e' \) matches \( w \) then, since \( |w| > 2|e'| \), by the induction hypothesis there exist \( x, y, z \) with \( |y| \geq 1 \) and \( |xy| \leq 2|e'| < n_0 \) such that \( e' \) (and therefore also \( e = e'|e'' \) matches \( xy^kz \) for every \( k \). The same arguments works in the case that \( e'' \) matches \( w \).

In the case (b), if \( w \) is matched by \( (e')(e'') \) then we can write \( w = w'w'' \) where \( e' \) matches \( w' \) and \( e'' \) matches \( w'' \). We split to subcases. If \( |w'| > 2|e'| \) then by the induction hypothesis there exist \( x, y, z \) with \( |y| \leq 1 \), \( |xy| \leq 2|e'| < n_0 \) such that \( w' = xyz' \) and \( e' \) matches \( xy^kz' \) for every \( k \in \mathbb{N} \). This completes the proof since if we set \( z = z'w'' \) then we see that \( w = w'w'' = xyz \) and \( e = (e')(e'')(e'') \) matches \( xy^kz \) for every \( k \in \mathbb{N} \). Otherwise, if \( |w'| \leq 2|e'| \) then since \( |w| = |w'| + |w''| > n_0 = 2(|e'| + |e''|) \), it must be that \( |w''| > 2|e''| \). Hence by the induction hypothesis there exist \( x', y, z \) such that \( |y| \geq 1 \), \( |x'y| \leq 2|e''| \) and \( e'' \) matches \( x'y^kz \) for every \( k \in \mathbb{N} \). But now if we set \( x = w'x' \) we see that \( |xy| \leq |w'| + |x'y| \leq 2|e'| + 2|e''| = n_0 \) and on the other hand the expression \( e = (e')(e'')(e'') \) matches \( xy^kz = w'x'y^kz \) for every \( k \in \mathbb{N} \).

In case (c), if \( w \) is matched \( (e')^* \) then \( w = w_0 \cdots w_1 \) where for every \( i \in [l], w_i \) is a nonempty string matched by \( e' \). If \( |w_0| > 2|e'| \) then we can use the same approach as in the concatenation case above. Otherwise, we simply note that if \( x \) is the empty string, \( y = w_0 \), and \( z = w_1 \cdots w_l \) then \( |xy| \leq n_0 \) and \( xy^kz \) is matched by \( (e')^* \) for every \( k \in \mathbb{N} \).
Remark 9.16 — Recursive definitions and inductive proofs. When an object is recursively defined (as in the case of regular expressions) then it is natural to prove properties of such objects by induction. That is, if we want to prove that all objects of this type have property $P$, then it is natural to use an inductive steps that says that if $o', o'', o'''$ etc have property $P$ then so is an object $o$ that is obtained by composing them.

Using the pumping lemma, we can easily prove Lemma 9.14 (i.e., the non-regularity of the “matching parenthesis” function):

Proof of Lemma 9.14. Suppose, towards the sake of contradiction, that there is an expression $e$ such that $\Phi_e = MATCHPAREN$. Let $n_0$ be the number from Lemma 9.14 and let $w = \langle n_0 \rangle^{n_0}$ (i.e., $n_0$ left parenthesis followed by $n_0$ right parenthesis). Then we see that if we write $w = xyz$ as in Lemma 9.14, the condition $|xy| \leq n_0$ implies that $y$ consists solely of left parenthesis. Hence the string $xy^2z$ will contain more left parenthesis than right parenthesis. Hence $MATCHPAREN(xy^2z) = 0$ but by the pumping lemma $\Phi_e(xy^2z) = 1$, contradicting our assumption that $\Phi_e = MATCHPAREN$.

The pumping lemma is a very useful tool to show that certain functions are not computable by a regular expression. However, it is not an “if and only if” condition for regularity: there are non regular functions that still satisfy the conditions of the pumping lemma. To understand the pumping lemma, it is important to follow the order of quantifiers in Theorem 9.15. In particular, the number $n_0$ in the statement of Theorem 9.15 depends on the regular expression (in the proof we chose $n_0$ to be twice the number of symbols in the expression). So, if we want to use the pumping lemma to rule out the existence of a regular expression $e$ computing some function $F$, we need to be able to choose an appropriate input $w \in \{0, 1\}^*$ that can be arbitrarily large and satisfies $F(w) = 1$. This makes sense if you think about the intuition behind the pumping lemma: we need $w$ to be large enough as to force the use of the star operator.

Solved Exercise 9.1 — Palindromes is not regular. Prove that the following function over the alphabet $\{0, 1, ;\}$ is not regular: $PAL(w) = 1$ if and only if $w = u; u^R$ where $u \in \{0, 1\}^*$ and $u^R$ denotes $u$ “reversed”: the string $u_{|u|-1} \cdots u_0$. (The Palindrome function is most often defined without an explicit separator character $;$, but the version with such a separator is a bit cleaner and so we use it here. This does not make much difference, as one can easily encode the separator as a special binary string instead.)
Exercise: Let $F: \{0,1\}^* \to \{0,1\}$ defined such that $F(x) = 1$ iff $x = 0^n1^n$ for $n \in \mathbb{N}$.
Prove that $F$ is not regular.

Blue Team: Student proving $F$ is not regular

Red Team: Hypothetical “adversary” claiming $F$ is regular

"$F$ is computed by a regular expression $exp$"

"Is that so? Then what is the number whose existence is guaranteed by the pumping lemma?"

"Here is the number – you can call it $n_0$"

"In this case, let me choose $w = 0^{n_0}1^{n_0}$. Notice that $F(w) = 1$. What is the partition $w = xyz$ from the pumping lemma?"

"Since $|xy| \leq n_0$ and $|y| \geq 1$, I guess I am forced to use $x = 0^a$, $y = 0^b$, $z = 0^{n_0-a-b}1^{n_0}$ for $b \geq 1$ and $a \leq n_0 - b$"

"In this case, since I can choose $k$ as I want, let me set $k = 2$ and note that $xy^2z = 0^{n_0+b}1^{n_0}$ which contradicts the pumping lemma conclusion that $F(xy^2z) = 1$!"

Pumping Lemma: If $exp$ computes $F$ there exists $n_0$ such that for every $w$ with $F(w) = 1$ and $|w| > n_0$ there exists partition $w = xyz$ with $|xy| \leq n_0$ and $|y| \geq 1$ such that for every $k \in \mathbb{N}$ it holds that $F(xy^kz) = 1$

Figure 9.7: A cartoon of a proof using the pumping lemma that a function $F$ is not regular. The pumping lemma states that if $F$ is regular then there exists a number $n_0$ such that for every large enough $w$ with $F(w) = 1$, there exists a partition of $w$ to $w = xyz$ satisfying certain conditions such that for every $k \in \mathbb{N}$, $F(xy^kz) = 1$. You can imagine a pumping-lemma based proof as a game between you and the adversary. Every quantifier corresponds to an object you are free to choose on your own (and base your choice on previously chosen objects). Every for every quantifier corresponds to an object the adversary can choose arbitrarily (and again based on prior choices) as long as it satisfies the conditions. A valid proof corresponds to a strategy by which no matter what the adversary does, you can win the game by obtaining a contradiction which would be a choice of $k$ that would result in $F(xy^kz) = 0$, hence violating the conclusion of the pumping lemma.
**Solution:**

We use the pumping lemma. Suppose towards the sake of contradiction that there is a regular expression $e$ computing $PAL$, and let $n_0$ be the number obtained by the pumping lemma (Theorem 9.15). Consider the string $w = 0^n; 0^n$. Since the reverse of the all zero string is the all zero string, $PAL(w) = 1$. Now, by the pumping lemma, if $PAL$ is computed by $e$, then we can write $w = xyz$ such that $|xy| \leq n_0, |y| \geq 1$ and $PAL(xy^kz) = 1$ for every $k \in \mathbb{N}$. In particular, it must hold that $PAL(xz) = 1$, but this is a contradiction, since $xz = 0^n a^{-|y|}; 0^n$ and so its two parts are not of the same length and in particular are not the reverse of one another.

For yet another example of a pumping-lemma based proof, see Fig. 9.7 which illustrates a cartoon of the proof of the non-regularity of the function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ which is defined as $F(x) = 1$ iff $x = 0^n1^n$ for some $n \in \mathbb{N}$ (i.e., $x$ consists of a string of consecutive zeroes, followed by a string of consecutive ones of the same length).

### 9.5 OTHER SEMANTIC PROPERTIES OF REGULAR EXPRESSIONS

Regular expressions are widely used beyond just searching. For example, regular expressions are often used to define *tokens* (such as what is a valid variable identifier, or keyword) in programming languages. But they also have other uses. One nice example is the recent work on the NetKAT network programming language. In recent years, the world of networking moved from fixed topologies to “software defined networks”. These are run by programmable switches that can implement policies such as “if packet is secured by SSL then forward it to A, otherwise forward it to B”. By its nature, one would want to use a formalism for such policies that is guaranteed to always halt (and quickly!) and such that it is possible to answer semantic questions such as “does C see the packets moved from A to B” etc. The NetKAT language uses a variant of regular expressions to achieve precisely that.

Such applications use the fact that because regular expressions are so restricted, we can not only solve the halting problem for them, but also answer other *semantic questions*. Such semantic questions would not be solvable for Turing-complete models due to Rice’s Theorem (Theorem 8.13). For example, we can tell whether two regular expressions are equivalent, as well as whether a regular expression computes the constant zero function.
Theorem 9.17 — Emptiness of regular languages is computable. There is an algorithm that given a regular expression $e$, outputs 1 if and only if $\Phi_e$ is the constant zero function.

Proof Idea:

The idea is that we can directly observe this from the structure of the expression. The only way a regular expression $e$ computes the constant zero function is if $e$ has the form $\emptyset$ or is obtained by concatenating $\emptyset$ with other expressions.

$\ast$

Proof of Theorem 9.17. Define a regular expression to be “empty” if it computes the constant zero function. Given a regular expression $e$, we can determine if $e$ is empty using the following rules:

- If $e$ has the form $\sigma$ or "$" then it is not empty.
- If $e$ is not empty then $e | e'$ is not empty for every $e'$.
- If $e$ is not empty then $e^*$ is not empty.
- If $e$ and $e'$ are both not empty then $e e'$ is not empty.
- $\emptyset$ is empty.

Using these rules it is straightforward to come up with a recursive algorithm to determine emptiness.

$\blacksquare$

Theorem 9.18 — Equivalence of regular expressions is computable. Let $REGEQ : \{0, 1\}^* \to \{0, 1\}$ be the function that on input (a string representing) a pair of regular expressions $e, e'$, $REGEQ(e, e') = 1$ if and only if $\Phi_e = \Phi_{e'}$. Then $REGEQ$ is computable.

Proof Idea:

The idea is to show that given a pair of regular expression $e$ and $e'$ we can find an expression $e''$ such that $\Phi_{e''}(x) = 1$ if and only if $\Phi_e(x) \neq \Phi_{e''}(x)$. Therefore $\Phi_{e''}$ is the constant zero function if and only if $e$ and $e'$ are equivalent, and thus we can test for emptiness of $e''$ to determine equivalence of $e$ and $e'$.

$\ast$

Proof of Theorem 9.18. We will prove Theorem 9.18 from Theorem 9.17. (The two theorems are in fact equivalent: it is easy to prove Theorem 9.17 from Theorem 9.18, since checking for emptiness is the same
as checking equivalence with the expression $\emptyset$.) Given two regular expressions $e$ and $e'$, we will compute an expression $e''$ such that $\Phi_{e''}(x) = 1$ if and only if $\Phi_e(x) \neq \Phi_{e'}(x)$. One can see that $e$ is equivalent to $e'$ if and only if $e''$ is empty.

We start with the observation that for every bits $a, b \in \{0, 1\}$, $a \neq b$ if and only if $(a \land b) \lor (\overline{a} \land b)$. Hence we need to construct $e''$ such that for every $x$,

$$\Phi_{e''}(x) = (\Phi_e(x) \land \overline{\Phi_{e'}(x)}) \lor (\overline{\Phi_e(x)} \land \Phi_{e'}(x)).$$

(9.10)

To construct the expression $e''$, we will show how given any pair of expressions $e$ and $e'$, we can construct expressions $e \land e'$ and $\overline{e}$ that compute the functions $\Phi_e \land \Phi_{e'}$ and $\overline{\Phi_e}$ respectively. (Computing the expression for $e \lor e'$ is straightforward using the $|$ operation of regular expressions.)

Specifically, by Lemma 9.13, regular functions are closed under negation, which means that for every regular expression $e$, there is an expression $\overline{e}$ such that $\Phi_{\overline{e}}(x) = 1 - \Phi_e(x)$ for every $x \in \{0, 1\}^*$. Now, for every two expression $e$ and $e'$, the expression

$$e \land e' = (\overline{e} \land \overline{e'})$$

(9.11)

computes the AND of the two expressions. Given these two transformations, we see that for every regular expressions $e$ and $e'$ we can find a regular expression $e''$ satisfying (9.10) such that $e''$ is empty if and only if $e$ and $e'$ are equivalent.

9.6 CONTEXT FREE GRAMMARS

If you have ever written a program, you’ve experienced a syntax error. You probably also had the experience of your program entering into an infinite loop. What is less likely is that the compiler or interpreter entered an infinite loop while trying to figure out if your program has a syntax error.

When a person designs a programming language, they need to determines its syntax. That is, the designer decides which strings corresponds to valid programs, and which ones do not (i.e., which strings contain a syntax error). To ensure that a compiler or interpreter always halts when checking for syntax errors, language designers typically do not use a general Turing-complete mechanism to express their syntax. Rather they use a restricted computational model. One of the most popular choices for such models is context free grammars.

To explain context free grammars, let us begin with a canonical example. Consider the function $ARITH : \Sigma^* \rightarrow \{0, 1\}$ that takes as input
a string \( x \) over the alphabet \( \Sigma = \{ (, ), +, −, \times, ÷, 0, 1, 2, 3, 4, 5, 6, 7, 8, 9 \} \) and returns 1 if and only if the string \( x \) represents a valid arithmetic expression. Intuitively, we build expressions by applying an operation such as \(+, −, \times, ÷\) to smaller expressions, or enclosing them in parenthesis, where the “base case” corresponds to expressions that are simply numbers. More precisely, we can make the following definitions:

- A **digit** is one of the symbols 0, 1, 2, 3, 4, 5, 6, 7, 8, 9.
- A **number** is a sequence of digits. (For simplicity we drop the condition that the sequence does not have a leading zero, though it is not hard to encode it in a context-free grammar as well.)
- An **operation** is one of \(+, −, \times, ÷\)
- An **expression** has either the form “number”, the form “sub-expression1 operation sub-expression2”, or the form “(sub-expression1)”, where “sub-expression1” and “sub-expression2” are themselves expressions. (Note that this is a recursive definition.)

A context free grammar (CFG) is a formal way of specifying such conditions. A CFG consists of a set of **rules** that tell us how to generate strings from smaller components. In the above example, one of the rules is “if \( \textit{exp1} \) and \( \textit{exp2} \) are valid expressions, then \( \textit{exp1} \times \textit{exp2} \) is also a valid expression”; we can also write this rule using the shorthand \( \textit{expression} \Rightarrow \textit{expression} \times \textit{expression} \). As in the above example, the rules of a context-free grammar are often recursive: the rule \( \textit{expression} \Rightarrow \textit{expression} \times \textit{expression} \) defines valid expressions in terms of itself. We now formally define context-free grammars:

**Definition 9.19 — Context Free Grammar.** Let \( \Sigma \) be some finite set. A context free grammar (CFG) over \( \Sigma \) is a triple \((V, R, s)\) such that:

- \( V \), known as the **variables**, is a set disjoint from \( \Sigma \).
- \( v \in V \) is known as the **initial variable**.
- \( R \) is a set of **rules**. Each rule is a pair \((v, z)\) with \( v \in V \) and \( z \in (\Sigma \cup V)^* \). We often write the rule \((v, z)\) as \( v \Rightarrow z \) and say that the string \( z \) **can be derived** from the variable \( v \).

**Example 9.20 — Context free grammar for arithmetic expressions.** The example above of well-formed arithmetic expressions can be captured formally by the following context free grammar:
• The alphabet $\Sigma$ is $\{(,)\, plus, \, minus, \, times, \, divide, \, 0, \, 1, \, 2, \, 3, \, 4, \, 5, \, 6, \, 7, \, 8, \, 9\}$

• The variables are $V = \{\text{expression}, \, \text{number}, \, \text{digit}, \, \text{operation}\}$.

• The rules are the set $R$ containing the following 19 rules:

  – The 4 rules $\text{operation} \Rightarrow +, \text{operation} \Rightarrow -, \text{operation} \Rightarrow \times$, and $\text{operation} \Rightarrow ÷$.
  – The 10 rules $\text{digit} \Rightarrow 0,..., \text{digit} \Rightarrow 9$.
  – The rule $\text{number} \Rightarrow \text{digit}$.
  – The rule $\text{number} \Rightarrow \text{digit} \, \text{number}$.
  – The rule $\text{expression} \Rightarrow \text{number}$.
  – The rule $\text{expression} \Rightarrow \text{expression} \, \text{operation} \, \text{expression}$.
  – The rule $\text{expression} \Rightarrow (\text{expression})$.

• The starting variable is $\text{expression}$.

People use many different notations to write context free grammars. One of the most common notations is the Backus–Naur form. In this notation we write a rule of the form $v \Rightarrow a$ (where $v$ is a variable and $a$ is a string) in the form $\langle v \rangle := a$. If we have several rules of the form $v \Rightarrow a, v \Rightarrow b, \text{and} v \Rightarrow c$ then we can combine them as $\langle v \rangle := a | b | c$. (In words we say that $v$ can derive either $a, b,$ or $c$.) For example, the Backus-Naur description for the context free grammar of Example 9.20 is the following (using ASCII equivalents for operations):

$$\begin{align*}
\text{operation} & := +|\cdot|\ast|/ \\
\text{digit} & := 0|1|2|3|4|5|6|7|8|9 \\
\text{number} & := \text{digit}|\text{digit} \, \text{number} \\
\text{expression} & := \text{number}|\text{expression} \, \text{operation} \, \text{expression} \\
& \rightarrow \text{expression}|(\text{expression})
\end{align*}$$

Another example of a context free grammar is the “matching parenthesis” grammar, which can be represented in Backus-Naur as follows:

$$\text{match} := ""|\text{match} \, \text{match}|(\text{match})$$

A string over the alphabet $\{\,(,)\}$ can be generated from this grammar (where $\text{match}$ is the starting expression and $""$ corresponds to the empty string) if and only if it consists of a matching set of parenthesis. In contrast, by Lemma 9.14 there is no regular expression that matches a string $x$ if and only if $x$ contains a valid sequence of matching parenthesis.
9.6.1 Context-free grammars as a computational model
We can think of a context-free grammar over the alphabet $\Sigma$ as defining a function that maps every string $x$ in $\Sigma^*$ to 1 or 0 depending on whether $x$ can be generated by the rules of the grammars. We now make this definition formally.

**Definition 9.21 — Deriving a string from a grammar.** If $G = (V, R, s)$ is a context-free grammar over $\Sigma$, then for two strings $\alpha, \beta \in (\Sigma \cup V)^*$ we say that $\beta$ can be derived in one step from $\alpha$, denoted by $\alpha \Rightarrow_G \beta$, if we can obtain $\beta$ from $\alpha$ by applying one of the rules of $G$. That is, we obtain $\beta$ by replacing in $\alpha$ one occurrence of the variable $v$ with the string $z$, where $v \Rightarrow z$ is a rule of $G$.

We say that $\beta$ can be derived from $\alpha$, denoted by $\alpha \Rightarrow^*_G \beta$, if it can be derived by some finite number $k$ of steps. That is, if there are $\alpha_1, \ldots, \alpha_{k-1} \in (\Sigma \cup V)^*$, so that $\alpha \Rightarrow_G \alpha_1 \Rightarrow_G \alpha_2 \Rightarrow_G \cdots \Rightarrow_G \alpha_{k-1} \Rightarrow_G \beta$.

We say that $x \in \Sigma^*$ is matched by $G = (V, R, s)$ if $x$ can be derived from the starting variable $s$ (i.e., if $s \Rightarrow^*_G x$). We define the function computed by $(V, R, s)$ to be the map $\Phi_{V,R,s} : \Sigma^* \rightarrow \{0, 1\}$ such that $\Phi_{V,R,s}(x) = 1$ iff $x$ is matched by $(V, R, s)$. A function $F : \Sigma^* \rightarrow \{0, 1\}$ is context free if $F = \Phi_{V,R,s}$ for some CFG $(V, R, s)$.

A priori it might not be clear that the map $\Phi_{V,R,s}$ is computable, but it turns out that this is the case.

**Theorem 9.22 — Context-free grammars always halt.** For every CFG $(V, R, s)$ over $\{0, 1\}$, the function $\Phi_{V,R,s} : \{0, 1\}^* \rightarrow \{0, 1\}$ is computable.

As usual we restrict attention to grammars over $\{0, 1\}$ although the proof extends to any finite alphabet $\Sigma$.

**Proof.** We only sketch the proof. We start with the observation we can convert every CFG to an equivalent version of Chomsky normal form, where all rules either have the form $u \rightarrow vw$ for variables $u, v, w$ or the form $u \rightarrow \sigma$ for a variable $u$ and symbol $\sigma \in \Sigma$, plus potentially the rule $s \rightarrow \varepsilon$ where $s$ is the starting variable.

The idea behind such a transformation is to simply add new variables as needed, and so for example we can translate a rule such as $v \rightarrow uvw$ into the three rules $v \rightarrow ur, r \rightarrow tw$ and $t \rightarrow \sigma$.

Using the Chomsky Normal form we get a natural recursive algorithm for computing whether $s \Rightarrow^*_G x$ for a given grammar $G$ and string $x$. We simply try all possible guesses for the first rule $s \rightarrow uv$ that is used in such a derivation, and then all possible ways to par-
tition \( x \) as a concatenation \( x = x'x'' \). If we guessed the rule and the partition correctly, then this reduces our task to checking whether \( u \Rightarrow_G x' \) and \( v \Rightarrow_G x'' \), which (as it involves shorter strings) can be done recursively. The base cases are when \( x \) is empty or a single symbol, and can be easily handled.

**Remark 9.23 — Parse trees.** While we focus on the task of deciding whether a CFG matches a string, the algorithm to compute \( \Phi_{V,R,s} \) actually gives more information than that. That is, on input a string \( x \), if \( \Phi_{V,R,s}(x) = 1 \) then the algorithm yields the sequence of rules that one can apply from the starting vertex \( s \) to obtain the final string \( x \). We can think of these rules as determining a tree with \( s \) being the root vertex and the sinks (or leaves) corresponding to the substrings of \( x \) that are obtained by the rules that do not have a variable in their second element. This tree is known as the parse tree of \( x \), and often yields very useful information about the structure of \( x \).

Often the first step in a compiler or interpreter for a programming language is a parser that transforms the source into the parse tree (also known as the abstract syntax tree). There are also tools that can automatically convert a description of a context-free grammars into a parser algorithm that computes the parse tree of a given string. (Indeed, the above recursive algorithm can be used to achieve this, but there are much more efficient versions, especially for grammars that have particular forms, and programming language designers often try to ensure their languages have these more efficient grammars.)

### 9.6.2 The power of context free grammars

Context free grammars can capture every regular expression:

**Theorem 9.24 — Context free grammars and regular expressions.** Let \( e \) be a regular expression over \( \{0, 1\} \), then there is a CFG \( (V, R, s) \) over \( \{0, 1\} \) such that \( \Phi_{V,R,s} = \Phi_e \).

**Proof.** We prove the theorem by induction on the length of \( e \). If \( e \) is an expression of one bit length, then \( e = 0 \) or \( e = 1 \), in which case we leave it to the reader to verify that there is a (trivial) CFG that computes it. Otherwise, we fall into one of the following case: case 1: \( e = e'e'' \), case 2: \( e = e'|e'' \) or case 3: \( e = (e')^* \) where in all cases \( e', e'' \) are shorter regular expressions. By the induction hypothesis have grammars \( (V', R', s') \) and \( (V'', R'', s'') \) that compute \( \Phi_{e'} \) and \( \Phi_{e''} \).
respectively. By renaming of variables, we can also assume without loss of generality that \( V' \) and \( V'' \) are disjoint.

In case 1, we can define the new grammar as follows: we add a new starting variable \( s \notin V \cup V' \) and the rule \( s \mapsto s's'' \). In case 2, we can define the new grammar as follows: we add a new starting variable \( s \notin V \cup V' \) and the rules \( s \mapsto s' \) and \( s \mapsto s'' \). Case 3 will be the only one that uses recursion. As before we add a new starting variable \( s \notin V \cup V' \), but now add the rules \( s \mapsto "" \) (i.e., the empty string) and also add, for every rule of the form \((s', \alpha) \in R'\), the rule \( s \mapsto s\alpha \) to \( R \).

We leave it to the reader as (a very good!) exercise to verify that in all three cases the grammars we produce capture the same function as the original expression.

It turns out that CFG’s are strictly more powerful than regular expressions. In particular, as we’ve seen, the “matching parenthesis” function \( \text{MATCHPAREN} \) can be computed by a context free grammar, whereas, as shown in Lemma 9.14, it cannot be computed by regular expressions. Here is another example:

**Solved Exercise 9.2 — Context free grammar for palindromes.** Let \( \text{PAL} : \{0, 1, ;\}^* \rightarrow \{0, 1\} \) be the function defined in Solved Exercise 9.1 where \( \text{PAL}(w) = 1 \) iff \( w \) has the form \( u; u^R \). Then \( \text{PAL} \) can be computed by a context-free grammar

**Solution:**

A simple grammar computing \( \text{PAL} \) can be described using Backus–Naur notation:

```
start ::= ; | 0 start 0 | 1 start 1
```

One can prove by induction that this grammar generates exactly the strings \( w \) such that \( \text{PAL}(w) = 1 \).

A more interesting example is computing the strings of the form \( w; v \) that are not palindromes:

**Solved Exercise 9.3 — Non palindromes.** Prove that there is a context free grammar that computes \( \text{NPAL} : \{0, 1, ;\}^* \rightarrow \{0, 1\} \) where \( \text{NPAL}(w) = 1 \) if \( w = u; v \) but \( v \neq u^R \).

**Solution:**

Using Backus–Naur notation we can describe such a grammar as follows
palindrome := ; | 0 palindrome 0 | 1 palindrome 1
different := 0 palindrome 1 | 1 palindrome 0
start := different | 0 start | 1 start | start
   ↪  0 | start 1

In words, this means that we can characterize a string $w$ such that $\text{NPAL}(w) = 1$ as having the following form

$$w = \alpha b u; u R' \beta$$  \hspace{1cm} (9.12)

where $\alpha, \beta, u$ are arbitrary strings and $b \neq b'$. Hence we can generate such a string by first generating a palindrome $u; u R$ (palindrome variable), then adding either 0 on the right and 1 on the left to get something that is not a palindrome (different variable), and then we can add arbitrary number of 0’s and 1’s on either end (the start variable).

### 9.6.3 Limitations of context-free grammars (optional)

Even though context-free grammars are more powerful than regular expressions, there are some simple languages that are not captured by context free grammars. One tool to show this is the context-free grammar analog of the “pumping lemma” (Theorem 9.15):

**Theorem 9.25 — Context-free pumping lemma.** Let $(V, R, s)$ be a CFG over $\Sigma$, then there is some numbers $n_0, n_1 \in \mathbb{N}$ such that for every $x \in \Sigma^*$ with $|x| > n_0$, if $\Phi_{V, R, s}(x) = 1$ then $x = abcde$ such that $|b| + |c| + |d| \leq n_1$, $|b| + |d| \geq 1$, and $\Phi_{V, R, s}(ab^k cd^k e) = 1$ for every $k \in \mathbb{N}$.

**Proof of Theorem 9.25.** We only sketch the proof. The idea is that if the total number of symbols in the rules of the grammar is $k_0$, then the only way to get $|x| > n_0$ with $\Phi_{V, R, s}(x) = 1$ is to use recursion. That is, there must be some variable $v \in V$ such that we are able to derive from $v$ the value $bvd$ for some strings $b, d \in \Sigma^*$, and then further on derive from $v$ some string $c \in \Sigma^*$ such that $bcd$ is a substring of $x$ (in other words, $x = abcde$ for some $a, e \in \{0, 1\}^*$). If we take the variable $v$ satisfying this requirement with a minimum number
of derivation steps, then we can ensure that $|bcd|$ is at most some constant depending on $n_0$ and we can set $n_1$ to be that constant $(n_1 = 10 \cdot |R| \cdot n_0)$ will do, since we will not need more than $|R|$ applications of rules, and each such application can grow the string by at most $n_0$ symbols).

Thus by the definition of the grammar, we can repeat the derivation to replace the substring $bcd$ in $x$ with $b^kcd^k$ for every $k \in \mathbb{N}$ while retaining the property that the output of $\Phi_{V,R,s}$ is still one. Since $bcd$ is a substring of $x$, we can write $x = abcd e$ and are guaranteed that $ab^kcd^k e$ is matched by the grammar for every $k$.

Using Theorem 9.25 one can show that even the simple function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ defined as follows:

$$F(x) = \begin{cases} 1 & x = ww \text{ for some } w \in \{0, 1\}^* \\ 0 & \text{otherwise} \end{cases}$$

is not context-free. (In contrast, the function $G : \{0, 1\}^* \rightarrow \{0, 1\}$ defined as $G(x) = 1$ iff $x = w_0 w_1 \cdots w_{n-1} w_{n-1} w_{n-2} \cdots w_0$ for some $w \in \{0, 1\}^*$ and $n = |w|$ is context free, can you see why?)

**Solved Exercise 9.4 — Equality is not context-free.** Let $EQ : \{0, 1, ;\}^* \rightarrow \{0, 1\}$ be the function such that $EQ(x) = 1$ if and only if $x = u;\ u$ for some $u \in \{0, 1\}^*$. Then $EQ$ is not context free.

**Solution:**

We use the context-free pumping lemma. Suppose towards the sake of contradiction that there is a grammar $G$ that computes $EQ$, and let $n_0$ be the constant obtained from Theorem 9.25.

Consider the string $x = 1^n 0^n ; 1^n 0^n$, and write it as $x = abcd e$ as per Theorem 9.25, with $|bcd| \leq n_0$ and with $|b| + |d| \geq 1$. By Theorem 9.25, it should hold that $EQ(ace) = 1$. However, by case analysis this can be shown to be a contradiction.

Firstly, unless $b$ is on the left side of the ; separator and $d$ is on the right side, dropping $b$ and $d$ will definitely make the two parts different. But if it is the case that $b$ is on the left side and $d$ is on the right side, then by the condition that $|bcd| \leq n_0$ we know that $b$ is a string of only zeros and $d$ is a string of only ones. If we drop $b$ and $d$ then since one of them is non empty, we get that there are either less zeroes on the left side than on the right side, or there are less ones on the right side than on the left side. In either case, we get that $EQ(ace) = 0$, obtaining the desired contradiction.
9.7 SEMANTIC PROPERTIES OF CONTEXT FREE LANGUAGES

As in the case of regular expressions, the limitations of context free grammars do provide some advantages. For example, emptiness of context free grammars is decidable:

**Theorem 9.26** — Emptiness for CFG's is decidable. There is an algorithm that on input a context-free grammar \( G \), outputs 1 if and only if \( \Phi_G \) is the constant zero function.

**Proof Idea:**

The proof is easier to see if we transform the grammar to Chomsky Normal Form as in Theorem 9.22. Given a grammar \( G \), we can recursively define a non-terminal variable \( v \) to be non empty if there is either a rule of the form \( v \Rightarrow \sigma \), or there is a rule of the form \( v \Rightarrow uw \) where both \( u \) and \( w \) are non empty. Then the grammar is non empty if and only if the starting variable \( s \) is non-empty.

*Proof of Theorem 9.26.* We assume that the grammar \( G \) in Chomsky Normal Form as in Theorem 9.22. We consider the following procedure for marking variables as “non empty”:

1. We start by marking all variables \( v \) that are involved in a rule of the form \( v \Rightarrow \sigma \) as non empty.
2. We then continue to mark \( v \) as non empty if it is involved in a rule of the form \( v \Rightarrow uw \) where \( u, w \) have been marked before.

We continue this way until we cannot mark any more variables. We then declare that the grammar is empty if and only if \( s \) has not been marked. To see why this is a valid algorithm, note that if a variable \( v \) has been marked as “non empty” then there is some string \( \alpha \in \Sigma^* \) that can be derived from \( v \). On the other hand, if \( v \) has not been marked, then every sequence of derivations from \( v \) will always have a variable that has not been replaced by alphabet symbols. Hence in particular \( \Phi_G \) is the all zero function if and only if the starting variable \( s \) is not marked “non empty”.

9.7.1 Uncomputability of context-free grammar equivalence (optional)

By analogy to regular expressions, one might have hoped to get an algorithm for deciding whether two given context free grammars are equivalent. Alas, no such luck. It turns out that the equivalence problem for context free grammars is uncomputable. This is a direct corollary of the following theorem:
Theorem 9.27 — Fullness of CFG’s is uncomputable. For every set \( \Sigma \), let \( CFGFULL_{\Sigma} \) be the function that on input a context-free grammar \( G \) over \( \Sigma \), outputs 1 if and only if \( G \) computes the constant 1 function. Then there is some finite \( \Sigma \) such that \( CFGFULL_{\Sigma} \) is uncomputable.

Theorem 9.27 immediately implies that equivalence for context-free grammars is uncomputable, since computing “fullness” of a grammar \( G \) over some alphabet \( \Sigma = \{ \sigma_0, \ldots, \sigma_{k-1} \} \) corresponds to checking whether \( G \) is equivalent to the grammar \( s \Rightarrow \cdot | s \sigma_0 | \cdots | s \sigma_{k-1} \). Note that Theorem 9.27 and Theorem 9.26 together imply that context-free grammars, unlike regular expressions, are not closed under complement. (Can you see why?) Since we can encode every element of \( \Sigma \) using \( \lceil \log |\Sigma| \rceil \) bits (and this finite encoding can be easily carried out within a grammar) Theorem 9.27 implies that fullness is also uncomputable for grammars over the binary alphabet.

Proof Idea:
We prove the theorem by reducing from the Halting problem. To do that we use the notion of configurations of NAND-TM programs, as defined in Definition 7.8. Recall that a configuration of a program \( P \) is a binary string \( s \) that encodes all the information about the program in the current iteration.

We define \( \Sigma \) to be \( \{0, 1\} \) plus some separator characters and define \( INVALID_P : \Sigma^* \rightarrow \{0, 1\} \) to be the function that maps every string \( L \in \Sigma^* \) to 1 if and only if \( L \) does not encode a sequence of configurations that correspond to a valid halting history of the computation of \( P \) on the empty input.

The heart of the proof is to show that \( INVALID_P \) is context-free. Once we do that, we see that \( P \) halts on the empty input if and only if \( INVALID_P(L) = 1 \) for every \( L \). To show that, we will encode the list in a special way that makes it amenable to deciding via a context-free grammar. Specifically we will reverse all the odd-numbered strings.

Proof of Theorem 9.27. We only sketch the proof. We will show that if we can compute \( CFGFULL \) then we can solve \( HALTONZERO \), which has been proven uncomputable in Theorem 8.7. Let \( M \) be an input Turing machine for \( HALTONZERO \). We will use the notion of configurations of a Turing machine, as defined in Definition 7.8.

Recall that a configuration of Turing machine \( M \) and input \( x \) captures the full state of \( M \) at some point of the computation. The particular details of configurations are not so important, but what you need to remember is that:

* A configuration can be encoded by a binary string \( \sigma \in \{0, 1\}^* \).
• The initial configuration of $M$ on the input $0$ is some fixed string.

• A halting configuration will have the value a certain state (which can be easily “read off” from it) set to 1.

• If $\sigma$ is a configuration at some step $i$ of the computation, we denote by $\text{NEXT}_M(\sigma)$ as the configuration at the next step. $\text{NEXT}_M(\sigma)$ is a string that agrees with $\sigma$ on all but a constant number of coordinates (those encoding the position corresponding to the head position and the two adjacent ones). On those coordinates, the value of $\text{NEXT}_M(\sigma)$ can be computed by some finite function.

We will let the alphabet $\Sigma = \{0, 1\} \cup \{|, \#\}$. A computation history of $M$ on the input $0$ is a string $L \in \Sigma$ that corresponds to a list $\|\sigma_0\#\sigma_1\|\sigma_2\#\sigma_3\ldots\|\sigma_{t-2}\#\|\sigma_{t-1}\#\ldots\#\sigma_t\|$ (i.e., $|$ comes before an even numbered block, and $|$ comes before an odd numbered one) such that if $i$ is even then $\sigma_i$ is the string encoding the configuration of $P$ on input $0$ at the beginning of its $i$-th iteration, and if $i$ is odd then it is the same except the string is reversed. (That is, for odd $i$, $\text{rev}(\sigma_i)$ encodes the configuration of $P$ on input $0$ at the beginning of its $i$-th iteration.) Reversing the odd-numbered blocks is a technical trick to ensure that the function $\text{INVALID}_M$ we define below is context free.

We now define $\text{INVALID}_M : \Sigma^* \rightarrow \{0, 1\}$ as follows:

$$\text{INVALID}_M(L) = \begin{cases} 0 & \text{if } L \text{ is a valid computation history of } M \text{ on } 0 \\ 1 & \text{otherwise} \end{cases} \quad (9.14)$$

We will show the following claim:

**CLAIM:** $\text{INVALID}_M$ is context-free.

The claim implies the theorem. Since $M$ halts on $0$ if and only if there exists a valid computation history, $\text{INVALID}_M$ is the constant one function if and only if $M$ does not halt on $0$. In particular, this allows us to reduce determining whether $M$ halts on $0$ to determining whether the grammar $G_M$ corresponding to $\text{INVALID}_M$ is full.

We now turn to the proof of the claim. We will not show all the details, but the main point $\text{INVALID}_M(L) = 1$ if at least one of the following three conditions hold:

1. $L$ is not of the right format, i.e. not of the form $\langle\text{binary-string}\rangle\#\langle\text{binary-string}\rangle\|\langle\text{binary-string}\rangle\#\ldots$.

2. $L$ contains a substring of the form $\|\sigma\#\sigma'\|$ such that $\sigma' \neq \text{rev}(\text{NEXT}_P(\sigma))$

3. $L$ contains a substring of the form $\#\sigma\sigma'\#$ such that $\sigma' \neq \text{NEXT}_P(\text{rev}(\sigma))$
Since context-free functions are closed under the OR operation, the claim will follow if we show that we can verify conditions 1, 2 and 3 via a context-free grammar.

For condition 1 this is very simple: checking that $L$ is of the correct format can be done using a regular expression. Since regular expressions are closed under negation, this means that checking that $L$ is not of this format can also be done by a regular expression and hence by a context-free grammar.

For conditions 2 and 3, this follows via very similar reasoning to that showing that the function $F$ such that $F(u\#v) = 1$ iff $u \neq rev(v)$ is context-free, see Solved Exercise 9.3. After all, the $NEXT_M$ function only modifies its input in a constant number of places. We leave filling out the details as an exercise to the reader. Since $INVALID_M(L) = 1$ if and only if $L$ satisfies one of the conditions 1., 2. or 3., and all three conditions can be tested for via a context-free grammar, this completes the proof of the claim and hence the theorem.

9.8 SUMMARY OF SEMANTIC PROPERTIES FOR REGULAR EXPRESSIONS AND CONTEXT-FREE GRAMMARS

To summarize, we can often trade expressiveness of the model for amenability to analysis. If we consider computational models that are not Turing complete, then we are sometimes able to bypass Rice’s Theorem and answer certain semantic questions about programs in such models. Here is a summary of some of what is known about semantic questions for the different models we have seen.

<table>
<thead>
<tr>
<th>Model</th>
<th>Halting</th>
<th>Emptiness</th>
<th>Equivalence</th>
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<tbody>
<tr>
<td>Regular expressions</td>
<td>Computable</td>
<td>Computable</td>
<td>Computable</td>
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<tr>
<td>Context-free grammars</td>
<td>Computable</td>
<td>Computable</td>
<td>Uncomputable</td>
</tr>
<tr>
<td>Turing-complete models</td>
<td>Uncomputable</td>
<td>Uncomputable</td>
<td>Uncomputable</td>
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</tbody>
</table>

Lecture Recap

- The uncomputability of the Halting problem for general models motivates the definition of restricted computational models.
- In some restricted models we can answer semantic questions such as: does a given program terminate, or do two programs compute the same function?
- Regular expressions are a restricted model of computation that is often useful to capture tasks of
string matching. We can test efficiently whether an expression matches a string, as well as answer questions such as Halting and Equivalence.

• Context free grammars is a stronger, yet still not Turing complete, model of computation. The halting problem for context free grammars is computable, but equivalence is not computable.

9.9 EXERCISES

Exercise 9.1 — Closure properties of regular functions. Suppose that $F, G : \{0, 1\}^* \rightarrow \{0, 1\}$ are regular. For each one of the following definitions of the function $H$, either prove that $H$ is always regular or give a counterexample for regular $F, G$ that would make $H$ not regular.

1. $H(x) = F(x) \lor G(x)$.
2. $H(x) = F(x) \land G(x)$
3. $H(x) = \text{NAND}(F(x), G(x))$.
4. $H(x) = F(x^R)$ where $x^R$ is the reverse of $x$: $x^R = x_{n-1}x_{n-2}\ldots x_o$ for $n = |x|$.
5. $H(x) = \begin{cases} 1 & x = uv \text{ s.t. } F(u) = G(v) = 1 \\ 0 & \text{otherwise} \end{cases}$
6. $H(x) = \begin{cases} 1 & x = uu \text{ s.t. } F(u) = G(u) = 1 \\ 0 & \text{otherwise} \end{cases}$
7. $H(x) = \begin{cases} 1 & x = uu^R \text{ s.t. } F(u) = G(u) = 1 \\ 0 & \text{otherwise} \end{cases}$

Exercise 9.2 One among the following two functions that map $\{0, 1\}^*$ to $\{0, 1\}$ can be computed by a regular expression, and the other one cannot. For the one that can be computed by a regular expression, write the expression that does it. For the one that cannot, prove that this cannot be done using the pumping lemma. * $F(x) = 1$ if 4 divides $\sum_{i=0}^{\lfloor |x|/2 \rfloor} x_i$ and $F(x) = 0$ otherwise.
• $G(x) = 1$ if and only if $\sum_{i=0}^{\lfloor |x|/2 \rfloor} x_i \geq |x|/4$ and $G(x) = 0$ otherwise.

Exercise 9.3 — Closure properties of context-free functions. Suppose that $F, G : \{0, 1\}^* \rightarrow \{0, 1\}$ are context free. For each one of the following definitions of the function $H$, either prove that $H$ is always context
free or give a counterexample for regular $F, G$ that would make $H$ not context free.

1. $H(x) = F(x) \lor G(x)$.
2. $H(x) = F(x) \land G(x)$
3. $H(x) = \text{NAND}(F(x), G(x))$.
4. $H(x) = F(x^R)$ where $x^R$ is the reverse of $x$: $x^R = x_{n-1}x_{n-2}\cdots x_o$ for $n = |x|$.
5. $H(x) = \begin{cases} 1 & x = uv \text{ s.t. } F(u) = G(v) = 1 \\ 0 & \text{otherwise} \end{cases}$
6. $H(x) = \begin{cases} 1 & x = uu \text{ s.t. } F(u) = G(u) = 1 \\ 0 & \text{otherwise} \end{cases}$
7. $H(x) = \begin{cases} 1 & x = uu^R \text{ s.t. } F(u) = G(u) = 1 \\ 0 & \text{otherwise} \end{cases}$

Exercise 9.4 Prove that the function $F : \{0, 1\}^* \to \{0, 1\}$ such that $F(x) = 1$ if and only if $|x|$ is a power of two is not context free.

Exercise 9.5 — Syntax for programming languages. Consider the following syntax of a “programming language” whose source can be written using the ASCII character set:

- **Variables** are obtained by a sequence of letters, numbers and underscores, but can’t start with a number.

- A **statement** has either the form `foo = bar;` where `foo` and `bar` are variables, or the form `IF (foo) BEGIN ... END` where ... is list of one or more statements, potentially separated by newlines.

A **program** in our language is simply a sequence of statements (possibly separated by newlines or spaces).

1. Let $\text{VAR} : \{0, 1\}^* \to \{0, 1\}$ be the function that given a string $x \in \{0, 1\}^*$, outputs 1 if and only if $x$ corresponds to an ASCII encoding of a valid variable identifier. Prove that $\text{VAR}$ is regular.
2. Let $\text{SYN} : \{0, 1\}^* \to \{0, 1\}$ be the function that given a string $s \in \{0, 1\}^*$, outputs 1 if and only if $s$ is an ASCII encoding of a valid program in our language. Prove that $\text{SYN}$ is context free. (You do not have to specify the full formal grammar for $\text{SYN}$, but you need to show that such a grammar exists.)
3. Prove that $\text{SYN}$ is not regular. See footnote for hint$^3$

### 9.10 BIBLIOGRAPHICAL NOTES

The relation of regular expressions with finite automata is a beautiful topic, on which we only touch upon in this text. It is covered more extensively in [Hopcroft; Sip97; Koz97]. These texts also discuss topics such as non deterministic finite automata (NFA) and the relation between context-free grammars and pushdown automata.

Our proof of Theorem 9.6 is closely related to the Myhill-Nerode Theorem. One direction of the Myhill-Nerode theorem can be stated as saying that if $\mathcal{E}$ is a regular expression then there is at most a finite number of strings $z_0, \ldots, z_{k-1}$ such that $\Phi_{\mathcal{E}[z_i]} \neq \Phi_{\mathcal{E}[z_j]}$ for every $0 \leq i \neq j < k$.

As in the case of regular expressions, there are many resources available that cover context-free grammars in great detail. Chapter 2 of [Sip97] contains many examples of context-free grammars and their properties. There are also websites such as Grammophone where you can input grammars, and see what strings they generate, as well as some of the properties that they satisfy.

The adjective “context free” is used for CFG’s because a rule of the form $v \rightarrow a$ means that we can always replace $v$ with the string $a$, no matter what is the context in which $v$ appears. More generally, we might want to consider cases where the replacement rules depend on the context. This gives rise to the notion of general (aka “Type 0”) grammars that allow rules of the form $a \Rightarrow b$ where both $a$ and $b$ are strings over $(V \cup \Sigma)^*$. The idea is that if, for example, we wanted to enforce the condition that we only apply some rule such as $v \rightarrow 0w1$ when $v$ is surrounded by three zeroes on both sides, then we could do so by adding a rule of the form $000v000 \rightarrow 0000w1000$ (and of course we can add much more general conditions). Alas, this generality comes at a cost - general grammars are Turing complete and hence their halting problem is uncomputable. That is, there is no algorithm $A$ that can determine for every general grammar $G$ and a string $x$, whether or not the grammar $G$ generates $x$.

The Chomsky Hierarchy is a hierarchy of grammars from the least restrictive (most powerful) Type 0 grammars, which correspond to recursively enumerable languages (see Definition 8.16) to the most restrictive Type 3 grammars, which correspond to regular languages. Context-free languages correspond to Type 2 grammars. Type 1 grammars are context sensitive grammars. These are more powerful than context-free grammars but still less powerful than Turing machines. In particular functions/languages corresponding to context-sensitive
grammars are always computable, and in fact can be computed by a linear bounded automaton which are non-deterministic algorithms that take $O(n)$ space. For this reason, the class of functions/languages corresponding to context-sensitive grammars is also known as the complexity class $\text{NSPACE} O(n)$; we discuss space-bounded complexity in Chapter 16. While Rice’s Theorem implies that we cannot compute any non-trivial semantic property of Type 0 grammars, the situation is more complex for other types of grammars: some semantic properties can be determined and some cannot, depending on the grammar’s place in the hierarchy.
Is every theorem provable?

“Take any definite unsolved problem, such as ... the existence of an infinite number of prime numbers of the form $2^n + 1$. However unapproachable these problems may seem to us and however helpless we stand before them, we have, nevertheless, the firm conviction that their solution must follow by a finite number of purely logical processes...”

“...This conviction of the solvability of every mathematical problem is a powerful incentive to the worker. We hear within us the perpetual call: There is the problem. Seek its solution. You can find it by pure reason, for in mathematics there is no ignorabimus.”, David Hilbert, 1900.

“The meaning of a statement is its method of verification.”, Moritz Schlick, 1938 (aka “The verification principle” of logical positivism)

The problems shown uncomputable in Chapter 8, while natural and important, still intimately involved NAND-TM programs or other computing mechanisms in their definitions. One could perhaps hope that as long as we steer clear of functions whose inputs are themselves programs, we can avoid the “curse of uncomputability”. Alas, we have no such luck.

In this chapter we will see an example of a natural and seemingly “computation free” problem that nevertheless turns out to be uncomputable: solving Diophantine equations. As a corollary, we will see one of the most striking results of 20th century mathematics: Gödel’s Incompleteness Theorem, which showed that there are some mathematical statements (in fact, in number theory) that are inherently unprovable. We will actually start with the latter result, and then show the former.
10.1 HILBERT’S PROGRAM AND GÖDEL’S INCOMPLETENESS THEOREM

“And what are these ...vanishing increments? They are neither finite quantities, nor quantities infinitely small, nor yet nothing. May we not call them the ghosts of departed quantities?”, George Berkeley, Bishop of Cloyne, 1734.

The 1700’s and 1800’s were a time of great discoveries in mathematics but also of several crises. The discovery of calculus by Newton and Leibnitz in the late 1600’s ushered a golden age of problem solving. Many longstanding challenges succumbed to the new tools that were discovered, and mathematicians got ever better at doing some truly impressive calculations. However, the rigorous foundations behind these calculations left much to be desired. Mathematicians manipulated infinitesimal quantities and infinite series cavalierly, and while most of the time they ended up with the correct results, there were a few strange examples (such as trying to calculate the value of the infinite series $1 - 1 + 1 - 1 + 1 + ...$) which seemed to give out different answers depending on the method of calculation. This led to a growing sense of unease in the foundations of the subject which was addressed in works of mathematicians such as Cauchy, Weierstrass, and Riemann, who eventually placed analysis on firmer foundations, giving rise to the $\epsilon$’s and $\delta$’s that students taking honors calculus grapple with to this day.

In the beginning of the 20th century, there was an effort to replicate this effort, in greater rigor, to all parts of mathematics. The hope was to show that all the true results of mathematics can be obtained by starting with a number of axioms, and deriving theorems from them using logical rules of inference. This effort was known as the Hilbert program, named after the influential mathematician David Hilbert.

Alas, it turns out the results we’ve seen dealt a devastating blow to this program, as was shown by Kurt Gödel in 1931:

**Theorem 10.1 — Gödel’s Incompleteness Theorem: informal version.** For every sound proof system $V$ for sufficiently rich mathematical statements, there is a mathematical statement that is true but is not provable in $V$.

10.1.1 Defining “Proof Systems”

Before proving Theorem 10.3, we need to define “proof systems” and even formally define the notion of a “mathematical statement”. In geometry and other areas of mathematics, proof systems are often defined by starting with some basic assumptions or axioms and then
deriving more statements by using inference rules such as the famous Modus Ponens, but what axioms shall we use? What rules? We will use an extremely general notion of proof systems, not even restricting ourselves to ones that have the form of axioms and inference.

**Mathematical statements.** At the highest level, a mathematical statement is simply a piece of text, which we can think of as a string \( x \in \{0, 1\}^\ast \). Mathematical statements contain assertions whose truth does not depend on any empirical fact, but rather only on properties of abstract objects. For example, the following is a mathematical statement:

\[
\text{“The number } 2,696,635,869,504,783,333,238,805,675,613, 588,278,597,832,162,617,892,474,670,798,113 \text{ is prime”}.
\]

Mathematical statements do not have to involve numbers. They can assert properties of any other mathematical object including sets, strings, functions, graphs and yes, even programs. Thus, another example of a mathematical statement is the following:

```
def f(n):
    if n==1: return 1
    return f(3*n+1) if n % 2 else f(n//2)
```

**Proof systems.** A proof for a statement \( x \in \{0, 1\}^\ast \) is another piece of text \( w \in \{0, 1\}^\ast \) that certifies the truth of the statement asserted in \( x \). The conditions for a valid proof system are:

1. (Effectiveness) Given a statement \( x \) and a proof \( w \), there is an algorithm to verify whether or not \( w \) is a valid proof for \( x \). (For example, by going line by line and checking that each line follows from the preceding ones using one of the allowed inference rules.)

2. (Soundness) If there is a valid proof \( w \) for \( x \) then \( x \) is true.

These are quite minimal requirements for a proof system. Requirement 2 (soundness) is the very definition of a proof system: you shouldn’t be able to prove things that are not true. Requirement 1 is also essential. If it there is no set of rules (i.e., an algorithm) to check that a proof is valid then in what sense is it a proof system? We could replace it with the system where the “proof” for a statement \( x \) is “trust me: it’s true”.

We formally define proof systems as an algorithm \( V \) where \( V(x, w) = 1 \) holds if the string \( w \) is a valid proof for the statement \( x \).
Even if $x$ is true, the string $w$ does not have to be a valid proof for it (there are plenty of wrong proofs for true statements such as $4=2+2$) but if $w$ is a valid proof for $x$ then $x$ must be true.

**Definition 10.2 — Proof systems.** Let $\mathcal{T} \subseteq \{0, 1\}^*$ be some set (which we consider the “true” statements). A proof system for $\mathcal{T}$ is an algorithm $V$ that satisfies:

1. (Effectiveness) For every $x, w \in \{0, 1\}^*$, $V(x, w)$ halts with an output of either 0 or 1.
2. (Soundness) For every $x \notin \mathcal{T}$ and $w \in \{0, 1\}^*$, $V(x, w) = 0$.

A true statement $x \in \text{mathcal{T}$ is unprovable (with respect to $V$) if for every $w \in \{0, 1\}^*$, $V(x, w) = 0$. We say that $V$ is complete if there does not exist a true statement $x$ that is unprovable with respect to $v$.

### 10.2 Gödel’s Incompleteness Theorem: Computational Variant

Our first formalization of Theorem 10.3 involves statements about Turing machines. We let $\mathcal{H}$ be the set of strings $x \in \{0, 1\}^*$ that have the form “Turing machine $M$ halts on the zero input”.

**Theorem 10.3 — Gödel’s Incompleteness Theorem: computational variant.**

There does not exist a complete proof system for $\mathcal{H}$.

**Proof Idea:**

If we had such a complete and sound proof system then we could solve the HALTONZERO problem. On input a Turing machine $M$, we would search all purported proofs $w$ and halt as soon as we find a proof of either “$M$ halts on zero” or “$M$ does not halt on zero”. If the system is sound and complete then we will eventually find such a proof, and it will provide us with the correct output.

**Proof of Theorem 10.3.** Assume for the sake of contradiction that there was such a proof system $V$. We will use $V$ to build an algorithm $A$ that computes HALTONZERO, hence contradicting Theorem 8.7. Our algorithm $A$ will will work as follows:
Algorithm 10.4 — Halting from proofs.

Algorithm $A$:

- **Input:** Turing Machine $M$
- **Goal:** Determine if $M$ halts on the input 0.
- **Assumption:** We have access to a proof system $V$ such that for every statement $x$ of the form "Machine $M$ halts on 0" or "Machine $M$ does not halt on 0", there exists some string $w \in \{0, 1\}^*$ such that $V(x, w) = 1$ if and only if $x$ is true.

**Operation:**

- For $n = 0, 1, 2, \ldots$:
  - For $w \in \{0, 1\}^n$:
    * If $V("M \text{ halts on } 0", w) = 1$ output 1
    * If $V("M \text{ does not halt on } 0", w) = 1$ output 0

If $M$ halts on 0 then under our assumption there exists $w$ that proves this fact, and so when Algorithm $A$ reaches $n = |w|$ we will eventually find this $w$ and output 1, unless we already halted before. But we cannot halt before and output a wrong answer because it would contradict the soundness of the proof system. Similarly, this shows that if $M$ does not halt on 0 then (since we assume there is a proof of this fact too) our algorithm $A$ will eventually halt and output 0.

\[\blacksquare\]

**Remark 10.5 — The Gödel statement (optional).** One can extract from the proof of Theorem 10.3 a procedure that for every proof system $V$, yields a true statement $x^*$ that cannot be proven in $V$. But Gödel’s proof gave a very explicit description of such a statement $x^*$ which is closely related to the "Liar’s paradox". That is, Gödel’s statement $x^*$ was designed to be true if and only if $\forall_{w \in \{0, 1\}^*} V(x, w) = 0$. In other words, it satisfied the following property

\[x^* \text{ is true } \iff x^* \text{ does not have a proof in } V \quad (10.1)\]

One can see that if $x^*$ is true, then it does not have a proof, but if it is false then (assuming the proof system is sound) then it cannot have a proof, and hence $x^*$ must be both true and unprovable. One might wonder
how is it possible to come up with an $x^*$ that satisfies a condition such as (10.1) where the same string $x^*$ appears on both the righthand side and the lefthand side of the equation. The idea is that the proof of Theorem 10.3 yields a way to transform every statement $x$ into a statement $F(x)$ that is true if and only if $x$ does not have a proof in $V$. Thus $x^*$ needs to be a fixed point of $F$: a sentence such that $x^* = F(x^*)$. It turns out that we can always find such a fixed point of $F$. We’ve already seen this phenomenon in the λ calculus, where the $Y$ combinator maps every $F$ into a fixed point $YF$ of $F$. This is very related to the idea of programs that can print their own code. Indeed, Scott Aaronson likes to describe Gödel’s statement as follows:

The following sentence repeated twice, the second time in quotes, is not provable in the formal system $V$. “The following sentence repeated twice, the second time in quotes, is not provable in the formal system $V.”

In the argument above we actually showed that $x^*$ is true, under the assumption that $V$ is sound. Since $x^*$ is true and does not have a proof in $V$, this means that we cannot carry the above argument in the system $V$, which means that $V$ cannot prove its own soundness (or even consistency: that there is no proof of both a statement and its negation). Using this idea, it’s not hard to get Gödel’s second incompleteness theorem, which says that every sufficiently rich $V$ cannot prove its own consistency. That is, if we formalize the statement $c^*$ that is true if and only if $V$ is consistent (i.e., $V$ cannot prove both a statement and the statement’s negation), then $c^*$ cannot be proven in $V$.

### 10.3 QUANTIFIED INTEGER STATEMENTS

<table>
<thead>
<tr>
<th>HALT</th>
<th>TILING</th>
<th>MixedQS</th>
<th>IntQS</th>
<th>Diophantine</th>
</tr>
</thead>
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<td>is uncomputable</td>
<td>is uncomputable</td>
<td>is uncomputable</td>
<td>is uncomputable</td>
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</tr>
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</table>

Figure 10.1: Outline of the steps in the proof of Gödel’s Incompleteness Theorem.
There is something “unsatisfying” about Theorem 10.3. Sure, it shows there are statements that are unprovable, but they don’t feel like “real” statements about math. After all, they talk about programs rather than numbers, matrices, or derivatives, or whatever it is they teach in math courses. It turns out that we can get an analogous result for statements such as “there are no positive integers \( x \) and \( y \) such that \( x^2 - 2 = y^7 \)”, or “there are positive integers \( x, y, z \) such that \( x^2 + y^6 = z^{11} \)” that only talk about natural numbers.\(^3\) It doesn’t get much more “real math” than this. Indeed, the 19th century mathematician Leopold Kronecker famously said that “God made the integers, all else is the work of man.”

To make this more precise, let us define the notion of quantified integer statements:

**Definition 10.6 — Quantified integer statements.** A quantified integer statement is a well-formed statement with no unbound variables involving integers, variables, the operators \( >, <, \times, +, -, = \), the logical operations \( \neg \) (NOT), \( \land \) (AND), and \( \lor \) (OR), as well as quantifiers of the form \( \exists_{x \in \mathbb{N}} \) and \( \forall_{y \in \mathbb{N}} \) where \( x, y \) are variable names.

We often care deeply about determining the truth of quantified integer statements. For example, the statement that Fermat’s Last Theorem is true for \( n = 3 \) can be phrased as the quantified integer statement

\[
\neg \exists_{a \in \mathbb{N}} \exists_{b \in \mathbb{N}} \exists_{c \in \mathbb{N}} (a > 0) \land (b > 0) \land (c > 0) \land (a \times a \times a + b \times b \times b = c \times c \times c).
\]

(10.2)

The twin prime conjecture, that states that there is an infinite number of numbers \( p \) such that both \( p \) and \( p + 2 \) are primes can be phrased as the quantified integer statement

\[
\forall_{n \in \mathbb{N}} \exists_{p \in \mathbb{N}} (p > n) \land \text{PRIME}(p) \land \text{PRIME}(p + 2)
\]

(10.3)

where we replace an instance of \( \text{PRIME}(q) \) with the statement \( (q > 1) \land \forall_{a \in \mathbb{N}} \forall_{b \in \mathbb{N}} (a = 1) \lor (a = q) \lor \neg (a \times b = q) \).

The claim (mentioned in Hilbert’s quote above) that are infinitely many primes of the form \( p = 2^n + 1 \) can be phrased as follows:

\[
\forall_{n \in \mathbb{N}} \exists_{p \in \mathbb{N}} (p > n) \land \text{PRIME}(p) \land \\
(\forall_{k \in \mathbb{N}} (k \neq 2 \land \text{PRIME}(k)) \Rightarrow \neg \text{DIVIDES}(k, p - 1))
\]

(10.4)

where \( \text{DIVIDES}(a, b) \) is the statement \( \exists_{c \in \mathbb{N}} b \times c = a \). In English, this corresponds to the claim that for every \( n \) there is some \( p > n \) such that all of \( p - 1 \)’s prime factors are equal to 2.

\(^3\) I do not know if these statements are actually true or false, see here.
Remark 10.7 — Syntactic sugar for quantified integer statements. To make our statements more readable, we often use syntactic sugar and so write \( x \neq y \) as shorthand for \( \neg(x = y) \), and so on. Similarly, the “implication operator” \( a \Rightarrow b \) is “syntactic sugar” or shorthand for \( \neg a \lor b \), and the “if and only if operator” \( a \iff b \) is shorthand for \( (a \Rightarrow b) \land (b \Rightarrow a) \). We will also allow ourselves the use of “macros”: plugging in one quantified integer statement in another, as we did with DIVIDES and PRIME above.

Much of number theory is concerned with determining the truth of quantified integer statements. Since our experience has been that, given enough time (which could sometimes be several centuries) humanity has managed to do so for the statements that it cared enough about, one could (as Hilbert did) hope that eventually we would be able to prove or disprove all such statements. Alas, this turns out to be impossible:

**Theorem 10.8 — Gödel’s Incompleteness Theorem for quantified integer statements.** Let \( V : \{0, 1\}^* \rightarrow \{0, 1\} \) a computable purported verification procedure for quantified integer statements. Then either:

- **\( V \) is not sound:** There exists a false statement \( x \) and a string \( w \in \{0, 1\}^* \) such that \( V(x, w) = 1 \).

  or

- **\( V \) is not complete:** There exists a true statement \( x \) such that for every \( w \in \{0, 1\}^* \), \( V(x, w) = 0 \).

**Theorem 10.8** is a direct corollary of the following result, just as **Theorem 10.3** was a direct corollary of the uncomputability of HALTONZERO:

**Theorem 10.9 — Uncomputability of quantified integer statements.** Let \( QIS : \{0, 1\}^* \rightarrow \{0, 1\} \) be the function that given a (string representation of) a quantified integer statement outputs 1 if it is true and 0 if it is false. Then \( QIS \) is uncomputable.

Since a quantified integer statement is simply a sequence of symbols, we can easily represent it as a string. For simplicity we will assume that every string represents some quantified integer statement, by mapping strings that do not correspond to such a statement to an arbitrary statement such as \( \exists x \in \mathbb{N} x = 1 \).
Please stop here and make sure you understand why the uncomputability of $QIS$ (i.e., Theorem 10.9) means that there is no sound and complete proof system for proving quantified integer statements (i.e., Theorem 10.8). This follows in the same way that Theorem 10.3 followed from the uncomputability of $HALTONZERO$, but working out the details is a great exercise (see Exercise 10.1).

In the rest of this chapter, we will show the proof of Theorem 10.8.

10.4 DIOPHANTINE EQUATIONS AND THE MRDP THEOREM

Many of the functions people wanted to compute over the years involved solving equations. These have a much longer history than mechanical computers. The Babylonians already knew how to solve some quadratic equations in 2000BC, and the formula for all quadratics appears in the Bakhshali Manuscript that was composed in India around the 3rd century. During the Renaissance, Italian mathematicians discovered generalization of these formulas for cubic and quartic (degrees 3 and 4) equations. Many of the greatest minds of the 17th and 18th century, including Euler, Lagrange, Leibniz and Gauss worked on the problem of finding such a formula for quintic equations to no avail, until in the 19th century Ruffini, Abel and Galois showed that no such formula exists, along the way giving birth to group theory.

However, the fact that there is no closed-form formula does not mean we can not solve such equations. People have been solving higher degree equations numerically for ages. The Chinese manuscript Jiuzhang Suanshu from the first century mentions such approaches. Solving polynomial equations is by no means restricted only to ancient history or to students’ homeworks. The gradient descent method is the workhorse powering many of the machine learning tools that have revolutionized Computer Science over the last several years.

But there are some equations that we simply do not know how to solve by any means. For example, it took more than 200 years until people succeeded in proving that the equation $a^{11} + b^{11} = c^{11}$ has no solution in integers. The notorious difficulty of so called Diophantine equations (i.e., finding integer roots of a polynomial) motivated the mathematician David Hilbert in 1900 to include the question of finding a general procedure for solving such equations in his famous list of twenty-three open problems for mathematics of the 20th century. I don’t think Hilbert doubted that such a procedure exists. After all, the whole history of mathematics up to this point involved the discovery of ever more powerful methods, and even impossibility results such

4 This is a special case of what’s known as “Fermat’s Last Theorem” which states that $a^n + b^n = c^n$ has no solution in integers for $n > 2$. This was conjectured in 1637 by Pierre de Fermat but only proven by Andrew Wiles in 1991. The case $n = 11$ (along with all other so called “regular prime exponents”) was established by Kummer in 1850.
as the inability to trisect an angle with a straightedge and compass, or the non-existence of an algebraic formula for quintic equations, merely pointed out to the need to use more general methods.

Alas, this turned out not to be the case for Diophantine equations. In 1970, Yuri Matiyasevich, building on a decades long line of work by Martin Davis, Hilary Putnam and Julia Robinson, showed that there is simply no method to solve such equations in general:

**Theorem 10.10 — MRDP Theorem.** Let $DIO : \{0, 1\}^* \rightarrow \{0, 1\}$ be the function that takes as input a string describing a 100-variable polynomial with integer coefficients $P(x_0, \ldots, x_{99})$ and outputs 1 if and only if there exists $z_0, \ldots, z_{99} \in \mathbb{N}$ s.t. $P(z_0, \ldots, z_{99}) = 0$.

Then $DIO$ is uncomputable.

As usual, we assume some standard way to express numbers and text as binary strings. The constant 100 is of course arbitrary; the problem is known to be uncomputable even for polynomials of degree four and at most 58 variables. In fact the number of variables can be reduced to nine, at the expense of the polynomial having a larger (but still constant) degree. See Jones’s paper for more about this issue.

**Remark 10.11 — Active code vs static data.** The difficulty in finding a way to distinguish between “code” such as NAND-TM programs, and “static content” such as polynomials is just another manifestation of the phenomenon that code is the same as data. While a fool-proof solution for distinguishing between the two is inherently impossible, finding heuristics that do a reasonable job keeps many firewall and anti-virus manufacturers very busy (and finding ways to bypass these tools keeps many hackers busy as well).

### 10.5 HARDNESS OF QUANTIFIED INTEGER STATEMENTS

We will not prove the MRDP Theorem (Theorem 10.10). However, as we mentioned, we will prove the uncomputability of $QIS$ (i.e., Theorem 10.9), which is a special case of the MRDP Theorem. The reason is that a Diophantine equation is a special case of a quantified integer statement where the only quantifier is $\exists$. This means that deciding the truth of quantified integer statements is a potentially harder problem than solving Diophantine equations, and so it is potentially easier to prove that $QIS$ is uncomputable.
If you find the last sentence confusing, it is worthwhile to reread it until you are sure you follow its logic. We are so accustomed to trying to find solutions for problems that it can sometimes be hard to follow the arguments for showing that problems are uncomputable.

Our proof of the uncomputability of QIS (i.e. Theorem 10.9) will, as usual, go by reduction from the Halting problem, but we will do so in two steps:

1. We will first use a reduction from the Halting problem to show that deciding the truth of quantified mixed statements is uncomputable. Quantified mixed statements involve both strings and integers. Since quantified mixed statements are a more general concept than quantified integer statements, it is easier to prove the uncomputability of deciding their truth.

2. We will then reduce the problem of quantified mixed statements to quantifier integer statements.

10.5.1 Step 1: Quantified mixed statements and computation histories
We define quantified mixed statements as statements involving not just integers and the usual arithmetic operators, but also string variables as well.

**Definition 10.12 — Quantified mixed statements.** A quantified mixed statement is a well-formed statement with no unbound variables involving integers, variables, the operators $>, <, \times, +, -, =$, the logical operations $\neg$ (NOT), $\land$ (AND), and $\lor$ (OR), as well as quantifiers of the form $\exists x \in \mathbb{N}, \exists y \in \{0, 1\}^*, \forall y \in \mathbb{N}, \forall b \in \{0, 1\}^*$ where $x, y, a, b$ are variable names. These also include the operator $|a|$ which returns the length of a string valued variable $a$, as well as the operator $a_i$ where $a$ is a string-valued variable and $i$ is an integer valued expression which is true if $i$ is smaller than the length of $a$ and the $i^{th}$ coordinate of $a$ is 1, and is false otherwise.

For example, the true statement that for every string $a$ there is a string $b$ that corresponds to $a$ in reverse order can be phrased as the following quantified mixed statement

$$
\forall a \in \{0, 1\}^*, \exists b \in \{0, 1\}^* (|a| = |b|) \land (\forall i \in \mathbb{N}, i < |a| \Rightarrow (a_i \Leftrightarrow b_{|a|-i})) .
$$

(10.5)

Quantified mixed statements are more general than quantified integer statements, and so the following theorem is potentially easier to prove than Theorem 10.9:
**Theorem 10.13 — Uncomputability of quantified mixed statements.** Let $QMS : \{0, 1\}^* \rightarrow \{0, 1\}$ be the function that given a (string representation of) a quantified mixed statement outputs $1$ if it is true and $0$ if it is false. Then $QMS$ is uncomputable.

**Proof Idea:**

The idea behind the proof is similar to that used in showing that one-dimensional cellular automata are Turing complete (Theorem 7.7) as well as showing that equivalence (or even “fullness”) of context free grammars is uncomputable (Theorem 9.27). We use the notion of a configuration of a NAND-TM program as in Definition 7.8. Such a configuration can be thought of as a string $\alpha$ over some large-but-finite alphabet $\Sigma$ describing its current state, including the values of all arrays, scalars, and the index variable $i$. It can be shown that if $\alpha$ is the configuration at a certain step of the execution and $\beta$ is the configuration at the next step, then $\beta_j = \alpha_j$ for all $j$ outside of $\{i - 1, i, i + 1\}$ where $i$ is the value of $i$. In particular, every value $\beta_j$ is simply a function of $\alpha_{j-1}, j, j+1$. Using these observations we can write a quantified mixed statement $\text{NEXT} (\alpha, \beta)$ that will be true if and only if $\beta$ is the configuration encoding the next step after $\alpha$. Since a program $P$ halts on input $x$ if and only if there is a sequence of configurations $\alpha^0, \ldots, \alpha^{t-1}$ (known as a computation history) starting with the initial configuration with input $x$ and ending in a halting configuration, we can define a quantified mixed statement to determine if there is such a statement by taking a universal quantifier over all strings $H$ (for history) that encode a tuple $(\alpha^0, \alpha^1, \ldots, \alpha^{t-1})$ and then checking that $\alpha^0$ and $\alpha^{t-1}$ are valid starting and halting configurations, and that $\text{NEXT} (\alpha^j, \alpha^{j+1})$ is true for every $j \in \{0, \ldots, t - 2\}$.

$\star$

**Proof of Theorem 10.13.** The proof will be obtained by a reduction from the Halting problem. Specifically, we will use the notion of a configuration of a Turing Machines (Definition 7.8) that we have seen in the context of proving that one dimensional cellular automata are Turing complete. We need the following facts about configurations:

- For every Turing Machine $P$, there is a finite alphabet $\Sigma$, and a configuration of $P$ is a string $\alpha \in \Sigma^*$.

- A configuration $\alpha$ encodes all the state of the program at a particular iteration, including the array, scalar, and index variables.

- If $\alpha$ is a configuration, then $\beta = \text{NEXT}_P (\alpha)$ denotes the configuration of the computation after one more iteration. $\beta$ is a string over $\Sigma$ of length either $|\alpha|$ or $|\alpha| + 1$, and every coordinate of $\beta$ is a function
of just three coordinates in $\alpha$. That is, for every $j \in \{0, \ldots, |\beta| - 1\}$, 
\[ \beta_j = \text{MAP}_P(\alpha_{j-1}, \alpha_j, \alpha_{j+1}) \]
where $\text{MAP}_P : \Sigma^3 \rightarrow \Sigma$ is some function depending on $P$.

- There are simple conditions to check whether a string $\alpha$ is a valid starting configuration corresponding to an input $x$, as well as to check whether a string $\alpha$ is an halting configuration. In particular these conditions can be phrased as quantified mixed statements.

- A program $P$ halts on input $x$ if and only if there exists a sequence of configurations $H = (\alpha_0, \alpha_1, \ldots, \alpha_{T-1})$ such that (i) $\alpha_0$ is a valid starting configuration of $P$ with input $x$, (ii) $\alpha_{T-1}$ is a valid halting configuration of $P$, and (iii) $\alpha_{i+1} = \text{NEXT}_P(\alpha_i)$ for every $i \in \{0, \ldots, T - 2\}$.

Let $U$ be a universal NAND-TM program. Such a program exists by Theorem 8.1. We define $\text{HALT}_U$ as the function such that $\text{HALT}_U(w) = 1$ if and only if $U$ halts on the input $w$. We claim that the function $\text{HALT}_U$ is uncomputable. Indeed, for every NAND-TM program $P$ (which we identify with its representation as a string) and input $x \in \{0, 1\}^*$ to $P$, $\text{HALT}(P, x) = \text{HALT}_U(\langle P, x \rangle)$ where $\langle P, x \rangle$ is some encoding of the pair $(P, x)$ as a string. Hence if we could compute $\text{HALT}_U$ then we could compute $\text{HALT}$, contradicting Theorem 8.4.

Let $\Sigma$ be the alphabet needed to encode configurations of $U$, and let $\ell = \lceil \log(|\Sigma| + 1) \rceil$. Then we can encode any symbol in $\Sigma \cup \{;\}$ (where “;” is some separator symbol we’ll use) as a string in $\{0, 1\}^\ell$, and so in particular can encode a sequence $\alpha^0; \alpha^1; \ldots; \alpha^T$ of configurations of $U$ as a single binary string which we’ll also name as $H$. Given any input $w \in \{0, 1\}^*$, we will create a mixed integer statement $\varphi_w$ that will have the following form:

\[
\varphi_w = \exists H \in \{0, 1\}^\ell \cdot H \text{ encodes a valid sequence of configurations of a halting computation of } U \text{ on } w
\]

(10.6)

The reasons we can encode this condition as an MIS are the following:

1. The conditions for checking that the initial configuration is valid are simple, and we can extract the first configuration from $H$ by first looking at an index $i$ which is a multiple of $\ell$ such that $H_i \ldots H_{i+\ell-1}$ encodes the separator symbol “;” and such that $i$ is the first such index. Another way to say it is that $i$ is the position of the first separator if there exists $k$ such that $i = k \times \ell$ and $H_{i \ldots i+\ell-1}$ and for every $j \in \mathbb{N}$, if $j < i$ then $H_{j \ldots j+\ell-1}$ does not encode “;”. This can be captured using the operators allowed in a quantified mixed statement and the $\forall$ and $\exists$ quantifiers.
2. We can similarly check that the last configuration is halting. Extracting the position \( i \) that encodes the last separator can be done in a way analogous to that of extracting the first one.

3. We can define a quantified mixed predicate \( \text{NEXT}(\alpha, \beta) \) that is true if and only if \( \beta = \text{NEXT}_U(\beta) \) (i.e., \( \beta \) encodes the configuration obtained by proceeding from \( \alpha \) in one computational step). Indeed \( \text{NEXT}(\alpha, \beta) \) is true if for every \( i \in \{0, \ldots, |\beta|\} \) which is a multiple of \( \ell \), \( \beta_{i,\ldots,i+\ell-1} = \text{MAP}_U(\alpha_{i-\ell,\ldots,i+2\ell-1}) \) where \( \text{MAP}_U : \{0, 1\}^{3\ell} \to \{0, 1\}^\ell \) is the finite function above (identifying elements of \( \Sigma \) with their encoding in \( \{0, 1\}^\ell \)). Since \( \text{MAP}_U \) is a finite function, we can express it using the logical operations AND, OR, NOT (for example by computing \( \text{MAP}_U \) with NAND’s).

4. We can then write the condition that for every substring of \( H \) that has the form \( \alpha \text{ENC}(;)\beta \) with \( \alpha, \beta \in \{0, 1\}^\ell \) and \( \text{ENC}(;) \) being the encoding of the separator “;”, it holds that \( \text{NEXT}(\alpha, \beta) \) is true.

Together the above yields a computable procedure that maps every input \( w \in \{0, 1\}^* \) to \( \text{HALT}_U \) into a quantified mixed statement \( \varphi_w \) such that \( \text{HALT}_U(w) = 1 \) if and only if \( \text{QMS}(\varphi_w) = 1 \). This reduces computing \( \text{HALT}_U \) to computing \( \text{QMS} \), and hence the uncomputability of \( \text{HALT}_U \) implies the uncomputability of \( \text{QMS} \).

10.5.2 Step 2: Reducing mixed statements to integer statements

We now show how to prove Theorem 10.9 using Theorem 10.13. The idea is again a proof by reduction. We will show a transformation of every quantifier mixed statement \( \varphi \) into a quantified integer statement \( \xi \) that does not use string-valued variables such that \( \varphi \) is true if and only if \( \xi \) is true.

To remove string-valued variables from a statement, we encode them by integers. We will show that we can encode a string \( x \in \{0, 1\}^* \) by a pair of numbers \( (X, n) \in \mathbb{N} \) s.t.

- \( n = |x| \)
- There is a quantified integer statement \( \text{COORD}(X, i) \) that for every \( i < n \), will be true if \( x_i = 1 \) and will be false otherwise.

This will mean that we can replace a “for all” quantifier over strings such as \( \forall x \in \{0, 1\}^* \) with a pair of quantifiers over integers of the form \( \forall X \in \mathbb{N} \forall n \in \mathbb{N} \) (and similarly replace an existential quantifier of the form \( \exists x \in \{0, 1\}^* \) with a pair of quantifiers \( \exists X \in \mathbb{N} \exists n \in \mathbb{N} \) ). We can later replace all calls to \( |x| \) by \( n \) and all calls to \( x_i \) by \( \text{COORD}(X, i) \). This means that if we are able to define \( \text{COORD} \) via a quantified integer statement, then we obtain a proof of Theorem 10.9, since we can use it to map
every mixed quantified statement \( \varphi \) to an equivalent quantified integer statement \( \xi \) such that \( \xi \) is true if and only if \( \varphi \) is true, and hence \( QMS(\varphi) = QIS(\xi) \). Such a procedure implies that the task of computing \( QMS \) reduces to the task of computing \( QIS \), which means that the uncomputability of \( QMS \) implies the uncomputability of \( QIS \).

The above shows that proof of the theorem all boils down to finding the right encoding of strings as integers, and the right way to implement \( COORD \) as a quantified integer statement. To achieve this we use the following technical result:

**Lemma 10.14 — Constructible prime sequence.** There is a sequence of prime numbers \( p_0 < p_1 < p_2 < \cdots \) such that there is a quantified integer statement \( PCOORD(p, i) \) that is true if and only if \( p = p_i \).

Using Lemma 10.14 we can encode a \( x \in \{0, 1\}^* \) by the numbers \((X, n)\) where \( X = \prod_{i=1}^{n} p_i \) and \( n = |x| \). We can then define the statement \( COORD(X, i) \) as

\[
\forall_{p \in \mathbb{N}} PCOORD(p, i) \Rightarrow DIVIDES(p, X)
\]  

(10.7)

where \( DIVIDES(a, b) \), as before, is defined as \( \exists_{c \in \mathbb{N}} a \times c = b \). Note that indeed if \( X, n \) encodes the string \( x \in \{0, 1\}^* \), then for every \( i < n \), \( COORD(X, i) = x_i \), since \( p_i \) divides \( X \) if and only if \( x_i = 1 \).

Thus all that is left to conclude the proof of Theorem 10.9 is to prove Lemma 10.14, which we now proceed to do.

**Proof.** The sequence of prime numbers we consider is the following: We fix \( C \) to be a sufficiently large constant (\( C = 2^{2^{34}} \) will do) and define \( p_i \) to be the smallest prime number that is in the interval \([i + C]^3 + 1, (i + C + 1)^3 - 1]\). It is known that there exists such a prime number for every \( i \in \mathbb{N} \). Given this, the definition of \( PCOORD(p, i) \) is simple:

\[
(p > (i+C) \times (i+C) \times (i+C)) \land (p < (i+C+1) \times (i+C+1) \times (i+C+1)) \land \left( \forall_{p'} \neg PRIME(p') \lor (p' \leq i) \lor (p' \geq p) \right)
\]  

(10.8)

We leave it to the reader to verify that \( PCOORD(p, i) \) is true iff \( p = p_i \). \( \blacksquare \)

To sum up we have shown that for every quantified mixed statement \( \varphi \), we can compute a quantified integer statement \( \xi \) such that \( QMS(\varphi) = 1 \) if and only if \( QIS(\xi) = 1 \). Hence the uncomputability of \( QMS \) (Theorem 10.13) implies the uncomputability of \( QIS \), completing the proof of Theorem 10.9, and so also the proof of Gödel’s Incompleteness Theorem for quantified integer statements (Theorem 10.8).
Lecture Recap

- Uncomputable functions include also functions that seem to have nothing to do with NAND-TM programs or other computational models such as determining the satisfiability of diophantine equations.
- This also implies that for any sound proof system (and in particular every finite axiomatic system) $S$, there are interesting statements $X$ (namely of the form “$F(x) = 0$” for an uncomputable function $F$) such that $S$ is not able to prove either $X$ or its negation.

10.6 Exercises

Exercise 10.1 — Gödel’s Theorem from uncomputability of $QIS$. Prove Theorem 10.8 using Theorem 10.9.

Exercise 10.2 — Expression for floor. Let $FSQRT(n, m) = \forall j \in \mathbb{N}((j \times j) > m) \lor (j \leq n)$. Prove that $FSQRT(n, m)$ is true if and only if $n = \lfloor \sqrt{m} \rfloor$.

Exercise 10.3 — Expression for computing the index. Recall that ?? asked you to prove that at iteration $t$ of a NAND-TM program the the variable $i$ is equal to $t - r(r + 1)$ if $t \leq (r + 1)^2$ and equals $(r + 2)(r + 1)t$ otherwise, where $r = \lfloor \sqrt{t + 1/4} - 1/2 \rfloor$. Prove that there is a quantified integer statement $INDEX$ with parameters $t, i$ such that $INDEX(t, i)$ is true if and $i$ is the value of $i$ after $t$ iterations.

Exercise 10.4 — Expression for computing the previous line. Give the following quantified integer expressions:

1. $MOD(a, b, c)$ which is true if and only if $b = a \mod c$. Note if a program has $s$ lines then the line executed at step $t$ is equal to $t \mod s$.

2. Suppose that $P$ is the three line NAND-CIRC program listed below. Give a quantified integer statement $LAST(n, t, t')$ such that $LAST(t, t')$ is true if and only if $t' - n$ is the largest step smaller than $t - n$ in which the variable on the righthand side of the line executed at step $t - n$ is written to. If this variable is an input variable $x_i$ then let $LAST(n, t, t')$ to be true if the current index location equals $t'$ and $t' < n$. 

y_0 := foo_i NAND foo_i
foo_i := x_i NAND x_i
loop := validx_i NAND validx_i

Exercise 10.5 — axiomatic proof systems. For every representation of logical statements as strings, we can define an axiomatic proof system to consist of a finite set of strings \( A \) and a finite set of rules \( I_0, \ldots, I_{m-1} \) with \( I_j : (\{0, 1\}^*)^k \rightarrow \{0, 1\}^* \) such that a proof \((s_1, \ldots, s_n)\) that \( s_n \) is true is valid if for every \( i \), either \( s_i \in A \) or is some \( j \in [m] \) and are \( i_1, \ldots, i_{k_j} < i \) such that \( s_i = I_j(s_{i_1}, \ldots, s_{i_{k_j}}) \). A system is sound if whenever there is no false \( s \) such that there is a proof that \( s \) is true. Prove that for every uncomputable function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) and every sound axiomatic proof system \( S \) (that is characterized by a finite number of axioms and inference rules), there is some input \( x \) for which the proof system \( S \) is not able to prove neither that \( F(x) = 0 \) nor that \( F(x) \neq 0 \).

Exercise 10.6 — MRDP exercise. The MRDP theorem states that the problem of determining, given a \( k \)-variable polynomial \( p \) with integer coefficients, whether there exists integers \( x_0, \ldots, x_{k-1} \) such that \( p(x_0, \ldots, x_{k-1}) = 0 \) is uncomputable. Consider the following quadratic integer equation problem: the input is a list of polynomials \( p_0, \ldots, p_{m-1} \) over \( k \) variables with integer coefficients, where each of the polynomials is of degree at most two (i.e., it is a quadratic function). The goal is to determine whether there exist integers \( x_0, \ldots, x_{k-1} \) that solve the equations \( p_0(x) = \cdots = p_{m-1}(x) = 0 \). Use the MRDP Theorem to prove that this problem is uncomputable. That is, show that the function \( QUADINTEQ : \{0, 1\}^* \rightarrow \{0, 1\} \) is uncomputable, where this function gets as input a string describing the polynomials \( p_0, \ldots, p_{m-1} \) (each with integer coefficients and degree at most two), and outputs 1 if and only if there exists \( x_0, \ldots, x_{k-1} \in \mathbb{Z} \) such that for every \( i \in [m] \), \( p_i(x_0, \ldots, x_{k-1}) = 0 \). See footnote for hint\(^5\)

Exercise 10.7 — The Busy Beaver problem. In this question we define the NAND-TM variant of the busy beaver function.

1. We define the function \( T : \{0, 1\}^* \rightarrow \mathbb{N} \) as follows: for every string \( P \in \{0, 1\}^* \), if \( P \) represents a NAND++ program such that when \( P \) is executed on the input 0 (i.e., the string of length 1 that is simply 0), a total of \( M \) lines are executed before the program halts, then \( T(P) = M \). Otherwise (if \( P \) does not represent a NAND++ program, or it is a program that does not halt on 0), \( T(P) = 0 \). Prove that \( T \) is uncomputable.

\(^5\) You can replace the equation \( y = x^4 \) with the pair of equations \( y = x^2 \) and \( z = x^2 \). Also, you can replace the equation \( w = yu, y = x^8 \) with the three equations \( w = yu, y = x^9 \) and \( u = x^2 \).
2. Let $TOWER(n)$ denote the number $2^{2^{2^{\cdots^{2}}}}$ (that is, a “tower of powers of two” of height $n$). To get a sense of how fast this function grows, $TOWER(1) = 2$, $TOWER(2) = 2^2 = 4$, $TOWER(3) = 2^{2^2} = 16$, $TOWER(4) = 2^{16} = 65536$ and $TOWER(5) = 2^{65536}$ which is about $10^{20000}$. $TOWER(6)$ is already a number that is too big to write even in scientific notation. Define $NBB : \mathbb{N} \to \mathbb{N}$ (for “NAND-TM Busy Beaver”) to be the function $NBB(n) = \max_{P \in \{0,1\}^n} T(P)$ where $T : \mathbb{N} \to \mathbb{N}$ is the function defined in Item 1. Prove that $NBB$ grows faster than $TOWER$, in the sense that $TOWER(n) = o(NBB(n))$ (i.e., for every $\epsilon > 0$, there exists $n_0$ such that for every $n > n_0$, $TOWER(n) < \epsilon \cdot NBB(n)$).\footnote{You will not need to use very specific properties of the $TOWER$ function in this exercise. For example, $NBB(n)$ also grows faster than the Ackerman function. You might find Aaronson’s blog post on the same topic to be quite interesting, and relevant to this book at large. If you like it then you might also enjoy this piece by Terence Tao.}

10.7 BIBLIOGRAPHICAL NOTES

As mentioned before, Gödel, Escher, Bach [Hof99] is a highly recommended book covering Gödel’s Theorem. A classic popular science book about Fermat’s Last Theorem is [Sin97].

Cantor’s are used for both Turing’s and Gödel’s theorems. In a twist of fate, using techniques originating from the works Gödel and Turing, Paul Cohen showed in 1963 that Cantor’s Continuum Hypothesis is independent of the axioms of set theory, which means that neither it nor its negation is provable from these axioms and hence in some sense can be considered as “neither true nor false” (see [Coh08]). The Continuum Hypothesis is the conjecture that for every subset $S$ of $\mathbb{R}$, either there is a one-to-one and onto map between $S$ and $\mathbb{N}$ or there is a one-to-one and onto map between $S$ and $\mathbb{R}$. It was conjectured by Cantor and listed by Hilbert in 1900 as one of the most important problems in mathematics. See also the non-conventional survey of Shelah [She03]. See here for recent progress on a related question.

Thanks to Alex Lombardi for pointing out an embarrassing mistake in the description of Fermat’s Last Theorem. (I said that it was open for exponent 11 before Wiles’ work.)
III

EFFICIENT ALGORITHMS
The problem of distinguishing prime numbers from composite and of resolving the latter into their prime factors is ... one of the most important and useful in arithmetic ... Nevertheless we must confess that all methods ... are either restricted to very special cases or are so laborious ... they try the patience of even the practiced calculator ... and do not apply at all to larger numbers.”, Carl Friedrich Gauss, 1798

“For practical purposes, the difference between algebraic and exponential order is often more crucial than the difference between finite and non-finite.”, Jack Edmunds, “Paths, Trees, and Flowers”, 1963

“What is the most efficient way to sort a million 32-bit integers?”, Eric Schmidt to Barack Obama, 2008
“I think the bubble sort would be the wrong way to go.”, Barack Obama.

So far we have been concerned with which functions are computable and which ones are not. In this chapter we look at the finer question of the time that it takes to compute functions, as a function of their input length. This is extremely important in the practice of computing. In introductory courses, coding interviews, and actual algorithm design, terms such as “$O(n)$ running time” are often used in an informal way. People don’t have a precise definition of what a linear-time algorithm is, but rather assume that “they’ll know it when they see it”. However, in this book we will make precise definitions, using our mathematical models of computation. This will allow us to ask (and sometimes answer) questions such as:

- “Is there a function that can be computed in $O(n^2)$ time but not in $O(n)$ time?”

Learning Objectives:
- Describe at a high level some interesting computational problems.
- The difference between polynomial and exponential time.
- Examples of techniques for obtaining efficient algorithms
- Examples of how seemingly small differences in problems can potentially make huge differences in their computational complexity.
• “Are there natural problems for which the best algorithm (and not just the best known) requires $2^\Omega(n)$ time?”

In this chapter we will survey some examples of computational problems, for some of which we know efficient (e.g., $n^c$-time for a small constant $c$) algorithms, and for others the best known algorithms are exponential. We want to get a feel as to the kinds of problems that lie on each side of this divide and also see how some seemingly minor changes in formulation can make the (known) complexity of a problem “jump” from polynomial to exponential. We will not formally define the notion of running time in this chapter, and so will use the same “I know it when I see it” notion of an $O(n)$ or $O(n^2)$ time algorithms as the one you’ve seen in introduction to computer science courses. In Chapter 12, we define running time precisely, using Turing Machines and the RAM machines / NAND-RAM formalisms.

While the difference between $O(n)$ and $O(n^2)$ time is often crucial for practice, our focus in this book will be on the difference between polynomial and exponential running time. One of the interesting phenomena of computing is that there is often a kind of a “threshold phenomenon” or “zero-one law” for running time. Many natural problems can either be solved in polynomial running time with a not-too-large exponent (e.g., something like $O(n^2)$ or $O(n^3)$), or require exponential (e.g., at least $2^{\Omega(n)}$ or $2^{\Omega(\sqrt{n})}$) time to solve. The reasons for this phenomenon are still not fully understood, but some light on this is shed by the concept of NP completeness, which we will see in Chapter 14. As we will see, questions about polynomial versus exponential time are often insensitive to the choice of the particular computational model, just as the question of whether a function $F$ is computable is insensitive to whether you use Turing machines, $\lambda$-calculus, or Javascript as your model of computation.

This chapter is merely a tiny sample of the landscape of computational problems and efficient algorithms. If you have not yet seen this material and want to explore it more deeply (which I very much hope you do!), the bibliographical notes contain references to some excellent texts, some of which are available freely on the web.

**Remark 11.1 — Relations between parts of this book.**

Part I of this book contained a quantitative study of computation of finite functions. We asked what are the resources (in terms of gates of Boolean circuits or lines in straight-line programs) required to compute various finite functions.

Part II of the book contained a qualitative study of computation of infinite functions (i.e., functions of unbounded input length). In that part we asked the
efficient computation

qualitative question of whether or not a function is computable at all, regardless of the number of operations. Part III of the book, beginning with this chapter, merges the two approaches and contains a quantitative study of computation of infinite functions. In this part we ask how do resources for computing a function scale with the length of the input. In Chapter 12 we define the notion of running time, and the class $P$ of functions that can be computed using a number of steps that scales polynomially with the input length. In Section 12.6 we will relate this class to the models of Boolean circuits and straightline programs that we studied in Part I.

11.1 PROBLEMS ON GRAPHS

In this chapter we discuss several examples of important computational problems. Many of the problems will involve graphs. We have already encountered graphs before (see Section 1.4.4) but let us now quickly recall the basic notation. A graph $G$ consists of a set of vertices $V$ and edges $E$ where each edge is a pair of vertices. In a directed graph, an edge is an ordered pair $(u, v)$, which we sometimes denote as $\overrightarrow{uv}$. In an undirected graph, an edge is an unordered pair (or simply a set) $\{u, v\}$ which we sometimes denote as $uv$ or $u \sim v$. An equivalent viewpoint is that an undirected graph is like a directed graph with the property that whenever the edge $\overrightarrow{uv}$ is present then so is the edge $\overrightarrow{vu}$.

In this chapter we will usually assume that the graphs are undirected and simple (i.e., containing no parallel edges or self-loops) unless stated otherwise. We denote by $n$ their number of vertices and $m$ the number of edges. To avoid annoying “edge cases” we will always assume that every vertex touches at least one edge, and so $m \geq n/2$. In fact, we typically think of the vertices as simply the set $[n]$ of the numbers from 0 till $n - 1$. Graphs can be represented either in the adjacency list or adjacency matrix representations. We can transform between these two representations using $O(n^2)$ operations, and hence for our purposes we will mostly consider them as equivalent.

There is a reason that graphs are so ubiquitous in computer science and other sciences. Graphs can be used to model a great many of the data that we encounter. These are not just the “obvious” data such as the road network (which can be thought of as a graph of whose vertices are locations with edges corresponding to road segments), or the web (which can be thought of as a graph whose vertices are web pages with edges corresponding to links), or social networks (which can be thought of as a graph whose vertices are people and the edges correspond to friend relation). Graphs can also denote correlations in data (e.g., graph of observations of features with edges corresponding
to features that tend to appear together), causal relations (e.g., gene regulatory networks, where a gene is connected to gene products it derives), or the state space of a system (e.g., graph of configurations of a physical system, with edges corresponding to states that can be reached from one another in one step).

11.1.1 Finding the shortest path in a graph

The shortest path problem is the task of, given a graph $G = (V, E)$ and two vertices $s, t \in V$, to find the length of the shortest path between $s$ and $t$ (if such a path exists). That is, we want to find the smallest number $k$ such that there are vertices $v_0, v_1, \ldots, v_k$ with $v_0 = s, v_k = t$ and for every $i \in \{0, \ldots, k - 1\}$ an edge between $v_i$ and $v_{i+1}$. Formally, we define $MINPATH : \{0, 1\}^* \rightarrow \{0, 1\}^*$ to be the function that on input a triple $(G, s, t)$ (represented as a string) outputs the number $k$ which is the length of the shortest path in $G$ between $s$ and $t$ or a string representing no path if no such path exists. (In practice people often want to also find the actual path itself and not just its length; it turns out that the algorithms to compute the length of the path often yield the actual path itself as a byproduct, and so everything we say about the task of computing the length also applies to the task of finding the path.)

If each vertex has at least two neighbors then there can be an exponential number of paths from $s$ to $t$, but fortunately we do not have to enumerate them all to find the shortest path. We can find the shortest path using a breadth first search (BFS), enumerating $s$’s neighbors, and then neighbors’ neighbors, etc., in order. If we maintain the neighbors in a list we can perform a BFS in $O(n^2)$ time, while using a queue we can do this in $O(m)$ time.\(^1\) Dijkstra’s algorithm is a well-known generalization of BFS to weighted graphs. More formally, the algorithm for computing the function $MINPATH$ is described in Algorithm 11.2.

\(^1\) A queue is a data structure for storing a list of elements in “First In First Out (FIFO)” order. Each “pop” operation removes an element from the queue in the order that they were “pushed” into it; see the Wikipedia page.
Algorithm 11.2 — Shortest path via BFS.

**Input:** Graph $G = (V, E)$ and vertices $s, t \in V$. Assume $V = [n]$.

**Output:** Length $k$ of shortest path from $s$ to $t$ or $\infty$ if no such path exists.

1. Let $D$ be length-$n$ array.
2. Set $D[s] = 0$ and $D[i] = \infty$ for all $i \in [n] \setminus \{s\}$.
3. Initialize queue $Q$ to contain $s$.
4. while $S$ non empty do
5. \hspace{1em} Pop $v$ from $Q$
6. \hspace{1em} if $v = t$ then
7. \hspace{1em} \hspace{1em} return $D[v]$
8. \hspace{1em} end if
9. \hspace{1em} for $u$ neighbor of $v$ with $D[u] = \infty$ do
10. \hspace{1em} \hspace{1em} Set $D[u] \leftarrow D[v] + 1$
11. \hspace{1em} \hspace{1em} Add $u$ to $Q$.
12. \hspace{1em} end for
13. end while
14. return $\infty$

Since we only add to the queue vertices $w$ with $D[w] = \infty$ (and then immediately set $D[w]$ to an actual number), we never push to the queue a vertex more than once, and hence the algorithm makes $n$ “push” and “pop” operations. For each vertex $v$, the number of times we run the inner loop is equal to the degree of $v$ and hence the total running time is proportional to the sum $2m$ of all degrees. Algorithm 11.2 returns the correct answer since the vertices are added to the queue in the order of their distance from $s$, and hence we will reach $t$ after we have explored all the vertices that are closer to $s$ than $t$.

**Remark 11.3 — On data structures.** If you’ve ever taken an algorithms course, you have probably encountered many data structures such as lists, arrays, queues, stacks, heaps, search trees, hash tables and many mores. Data structures are extremely important in computer science, and each one of those offers different tradeoffs between overhead in storage, operations supported, cost in time for each operation, and more. For example, if we store $n$ items in a list, we will need a linear (i.e., $O(n)$ time) scan to retrieve one of them, while we achieve the same operation in $O(1)$ time if we used a hash table. However, when we only care about polynomial-time algorithms, such factors of $O(n)$ in the running time will not make much differ-
11.1.2 Finding the longest path in a graph

The *longest path problem* is the task of, given a graph $G = (V, E)$ and two vertices $s, t \in V$, to find the length of the longest simple (i.e., non-intersecting) path between $s$ and $t$. If the graph is a road network, then the longest path might seem less motivated than the shortest path (unless you are the kind of person that always prefers the “scenic route”). But of graphs can and are used to model a variety of phenomena, and in many such cases finding the longest path (and some of its variants) can be very useful. In particular, finding the longest path is a generalization of the famous Hamiltonian path problem which asks for a maximally long simple path (i.e., path that visits all $n$ vertices once) between $s$ and $t$, as well as the notorious traveling salesman problem (TSP) of finding (in a weighted graph) a path visiting all vertices of cost at most $w$. TSP is a classical optimization problem, with applications ranging from planning and logistics to DNA sequencing and astronomy.

Surprisingly, while we can find the shortest path in $O(m)$ time, there is no known algorithm for the longest path problem that significantly improves on the trivial “exhaustive search” or “brute force” algorithm that enumerates all the exponentially many possibilities. Specifically the number of possible paths in an $n$ vertex graph is $\Theta(n^n)$. The best known algorithms for the longest path problem improve on this, but not by much: it takes $\Omega(c^n)$ time for some constant $c > 1$. (At the moment the best record is $c \sim 1.65$ or so; even obtaining an $O(2^n)$ time bound is not that simple, see Exercise 11.1.)

11.1.3 Finding the minimum cut in a graph

Given a graph $G = (V, E)$, a *cut* is a subset $S$ of $V$ such that $S$ is neither empty nor is it all of $V$. The edges cut by $S$ are those edges where one of their endpoints is in $S$ and the other is in $\overline{S} = V \setminus S$. We denote this set of edges by $E(S, \overline{S})$. If $s, t \in V$ then an $s, t$ cut is a cut such that $s \in S$ and $t \in \overline{S}$ (see Fig. 11.3). The *minimum $s, t$ cut problem* is the task of finding, given $s$ and $t$, the minimum number $k$ such that there is an $s, t$ cut cutting $k$ edges (the problem is also sometimes phrased as finding the set that achieves this minimum; it turns out that algorithms to compute the number often yield the set as well).

![Image](image.png)

*Figure 11.2*: A knight’s tour can be thought of as a maximally long path on the graph corresponding to a chessboard where we put an edge between any two squares that can be reached by one step via a legal knight move.
Formally, we define $\text{MINCUT} : \{0, 1\}^* \rightarrow \{0, 1\}^*$ to be the function that on input a string representing a triple $(G = (V, E), s, t)$ of a graph and two vertices, outputs the minimum number $k$ such that there exists a set $S \subseteq V$ with $s \in S$, $t \notin S$ and $|E(S, \overline{S})| = k$.

The minimum $s, t$ cut problem appears in many applications since minimum cuts often correspond to bottlenecks. For example, in a communication or railroad network the minimum cut between $s$ and $t$ corresponds to the smallest number of edges that, if dropped, will disconnect $s$ from $t$. (This was actually the original motivation for this problem; see Section 11.6.) Similar applications arise in scheduling and planning. In the setting of image segmentation, one can define a graph whose vertices are pixels and whose edges correspond to neighboring pixels of distinct colors. If we want to separate the foreground from the background then we can pick (or guess) a foreground pixel $s$ and background pixel $t$ and ask for a minimum cut between them.

The naive algorithm for computing $\text{MINCUT}$ will check all $2^n$ possible subsets of an $n$-vertex graph, but it turns out we can do much better than that. As we’ve seen in this book time and again, there is more than one algorithm to compute the same function, and some of those algorithms might be more efficient than others. Luckily this is one of those cases. There are faster algorithms that compute $\text{MINCUT}$ in time which is polynomial in the number of vertices (denoted by $\text{poly}(n)$).

**Min-Cut Max-Flow.** We can obtain a polynomial-time algorithm for computing $\text{MINCUT}$ using the Max-Flow Min-Cut Theorem. This theory says that the minimum cut between $s$ and $t$ equals the maximum amount of flow we can send from $s$ to $t$, if every edge has unit capacity. Specifically, imagine that every edge of the graph corresponded to a pipe that could carry one unit of fluid per one unit of time (say 1 liter of water per second). Now suppose we want to send a maximum amount of water per second from our source $s$ to the sink $t$. If there is an $s, t$-cut of at most $k$ edges, then this maximum will be at most $k$. Indeed, such a cut $S$ will be a “bottleneck” since at most $k$ units can flow from $S$ to its complement $\overline{S}$. The above reasoning can be used to show that the maximum flow from $s$ to $t$ is at most the value of the minimum $s, t$-cut. The surprising and non-trivial content of the Max-Flow Min-Cut Theorem is that the maximum flow is also at least the value of the minimum cut, and hence computing the cut is the same as computing the flow.

**11.1.4 Linear programming**

The Max-Flow Min-Cut Theorem reduces the task of computing a minimum cut to the task of computing a maximum flow. However, this still does not show how to compute such a flow. The Ford-Fulkerson
Algorithm is a direct way to compute a flow using incremental improvements. But computing flows in polynomial time is also a special case of a much more general tool known as **linear programming**.

A flow on a graph $G$ of $m$ edges can be modeled as a vector $x \in \mathbb{R}^m$ where for every edge $e$, $x_e$ corresponds to the amount of water per time-unit that flows on $e$. We think of an edge $e$ as an ordered pair $(u, v)$ (we can choose the order arbitrarily) and let $x_e$ be the amount of flow that goes from $u$ to $v$. (If the flow is in the other direction then we make $x_e$ negative.) Since every edge has capacity one, we know that $-1 \leq x_e \leq 1$ for every edge $e$. A valid flow has the property that the amount of water leaving the source $s$ is the same as the amount entering the sink $t$, and that for every other vertex $v$, the amount of water entering and leaving $v$ is the same.

Mathematically, we can write these conditions as follows:

$$
\sum_{e \ni s} x_e + \sum_{e \ni t} x_e = 0
$$

$$
\sum_{e \ni v} x_e = 0 \quad \forall \, v \in V \setminus \{s, t\}
$$

$$
-1 \leq x_e \leq 1 \quad \forall \, e \in E
$$

where for every vertex $v$, summing over $e \ni v$ means summing over all the edges that touch $v$.

The maximum flow problem can be thought of as the task of maximizing $\sum_{e \ni s} x_e$ over all the vectors $x \in \mathbb{R}^m$ that satisfy the above conditions (11.1). Maximizing a linear function $\ell(x)$ over the set of $x \in \mathbb{R}^m$ that satisfy certain linear equalities and inequalities is known as **linear programming**. Luckily, there are polynomial-time algorithms for solving linear programming, and hence we can solve the maximum flow (and so, equivalently, minimum cut) problem in polynomial time. In fact, there are much better algorithms for maximum-flow/minimum-cut, even for weighted directed graphs, with currently the record standing at $O(\min\{m^{10/7}, m^{1/2}\})$ time.

**Global minimum cut.** We can also define the problem of finding the **global minimum cut** (i.e., the non-empty and non-everything set $S$ that minimizes the number of edges cut). A polynomial time algorithm for the minimum $s, t$ cut can be used to solve the global minimum cut in polynomial time by enumerating over all pairs $s, t$.

### 11.1.5 Finding the maximum cut in a graph

The **maximum cut** problem is the task of finding, given an input graph $G = (V, E)$, the subset $S \subseteq V$ that maximizes the number of edges cut by $S$. (We could also define an $s, t$-cut variant of the maximum cut like we did for minimum cut; the two variants have similar complexity but we use define maximum cut here in the way that is more common
in the literature.) Like its cousin the minimum cut problem, the maximum cut problem is also very well motivated. For example, it arises in VLSI design, and also has some surprising relation to analyzing the Ising model in statistical physics.

Surprisingly, while there is a polynomial-time algorithm for the minimum cut problem, there is no known algorithm solving maximum cut much faster than the trivial “brute force” algorithm that tries all $2^n$ possibilities for the set $S$.

### 11.1.6 A note on convexity

There is an underlying reason for the sometimes radical difference between the difficulty of maximizing and minimizing a function over a domain. If $D \subseteq \mathbb{R}^n$, then a function $f : D \to \mathbb{R}$ is convex if for every $x, y \in D$ and $p \in [0, 1]$ \( f(px + (1-p)y) \leq pf(x) + (1-p)f(y) \). That is, if applied to the $p$-weighted midpoint between $x$ and $y$ is smaller than the $p$-weighted average value of $f$. If $D$ itself is convex (which means that if $x, y$ are in $D$ then so is the line segment between them), then this means that if $x$ is a local minimum of $f$ then it is also a global minimum. The reason is that if $f(y) < f(x)$ then every point $z = px + (1-p)y$ on the line segment between $x$ and $y$ will satisfy $f(z) \leq pf(x) + (1-p)f(y) < f(x)$ and hence in particular $x$ cannot be a local minimum. Intuitively, local minima of functions are much easier to find than global ones: after all, any “local search” algorithm that keeps finding a nearby point on which the value is lower, will eventually arrive at a local minima. One example of such a local search algorithm is gradient descent which takes a sequence of small steps, each one in the direction that would reduce the value by the most amount based on the current derivative.

Indeed, under certain technical conditions, we can often efficiently find the minimum of convex functions over a convex domain, and this underlies the reason problems such as minimum cut and shortest path are easy to solve. On the other hand, maximizing a convex function over a convex domain (or equivalently, minimizing a concave function) can often be a hard computational task. A linear function is both convex and concave, which is the reason both the maximization and minimization problems for linear functions can be done efficiently.

The minimum cut problem is not a priori a convex minimization task, because the set of potential cuts is discrete and not continuous. However, it turns out that we can embed it in a continuous and convex set via the (linear) maximum flow problem. The “max flow min cut” theorem ensuring that this embedding is “tight” in the sense that the minimum “fractional cut” that we obtain through the maximum-flow linear program will be the same as the true minimum cut. Un-
fortunately, we don’t know of such a tight embedding in the setting of the maximum cut problem.

The issue of convexity arises time and again in the context of computation. For example, one of the basic tasks in machine learning is empirical risk minimization. This is the task of finding a classifier for a set training examples. That is, the input is a list of labeled examples \((x_{m-1}, y_{m-1}), \ldots, (x_{m-1}, y_{m-1})\), where each \(x_i \in \{0, 1\}^n\) and \(y_i \in \{0, 1\}\), and the goal is to find a classifier \(h : \{0, 1\}^n \rightarrow \{0, 1\}\) (or sometimes \(h : \{0, 1\}^n \rightarrow \mathbb{R}\) that minimizes the number of errors. More generally, we want to find \(h\) that minimizes

\[
\sum_{i=0}^{m-1} L(y_i, h(x_i))
\]  

where \(L\) is some loss function. When \(L\) is the square loss function \(L(y, y') = (y - y')^2\) and \(h\) is a linear function, this corresponds to the well-known convex minimization task of linear regression. In other cases, when the task is non convex, there can be many global or local minima. That said, even if we don’t find the global (or even a local) minima, this continuous embedding can still help us. In particular, when running a local improvement algorithm such as Gradient Descent, we might still find a function \(h\) that is “useful” in the sense of having a small error on future examples from the same distribution.

11.2 BEYOND GRAPHS

Not all computational problems arise from graphs. We now list some other examples of computational problems that are of great interest.

11.2.1 SAT

A propositional formula \(\varphi\) involves \(n\) variables \(x_1, \ldots, x_n\) and the logical operators AND (\(\land\)), OR (\(\lor\)), and NOT (\(\neg\), also denoted as \(\cdot\)). We say that such a formula is in conjunctive normal form (CNF for short) if it is an AND of ORs of variables or their negations (we call a term of the form \(x_i\) or \(\overline{x_i}\) a literal). For example, this is a CNF formula

\[
(x_7 \lor x_{22} \lor x_{15}) \land (x_{37} \lor x_{22}) \land (x_{55} \lor \overline{x}_7)
\]  

The satisfiability problem is the task of determining, given a CNF formula \(\varphi\), whether or not there exists a satisfying assignment for \(\varphi\). A satisfying assignment for \(\varphi\) is a string \(x \in \{0, 1\}^n\) such that if \(\varphi\) evaluates to True if we assign its variables the values of \(x\). The SAT problem might seem as an abstract question of interest only in logic but in fact SAT is of huge interest in industrial optimization, with applications including manufacturing planning, circuit synthesis, software verification, air-traffic control, scheduling sports tournaments, and more.
2SAT. We say that a formula is a $k$-CNF if it is an AND of ORs where each OR involves exactly $k$ literals. The $k$-SAT problem is the restriction of the satisfiability problem for the case that the input formula is a $k$-CNF. In particular, the 2SAT problem is to find out, given a 2-CNF formula $\varphi$, whether there is an assignment $x \in \{0, 1\}^n$ that satisfies $\varphi$, in the sense that it makes it evaluate to 1 or “True”. The trivial, brute-force, algorithm for 2SAT will enumerate all the $2^n$ assignments $x \in \{0, 1\}^n$ but fortunately we can do much better. The key is that we can think of every constraint of the form $\ell_i \vee \ell_j$ (where $\ell_i, \ell_j$ are literals, corresponding to variables or their negations) as an implication $\overline{\ell_i} \Rightarrow \ell_j$, since it corresponds to the constraints that if the literal $\ell'_i = \overline{\ell_i}$ is true then it must be the case that $\ell_j$ is true as well. Hence we can think of $\varphi$ as a directed graph between the $2n$ literals, with an edge from $\ell_i$ to $\ell_j$ corresponding to an implication from the former to the latter. It can be shown that $\varphi$ is unsatisfiable if and only if there is a variable $x_i$ such that there is a directed path from $x_i$ to $\overline{x_i}$ as well as a directed path from $\overline{x_i}$ to $x_i$ (see Exercise 11.2). This reduces 2SAT to the (efficiently solvable) problem of determining connectivity in directed graphs.

3SAT. The 3SAT problem is the task of determining satisfiability for 3CNFs. One might think that changing from two to three would not make that much of a difference for complexity. One would be wrong. Despite much effort, we do not know of a significantly better than brute force algorithm for 3SAT (the best known algorithms take roughly $1.3^n$ steps).

Interestingly, a similar issue arises time and again in computation, where the difference between two and three often corresponds to the difference between tractable and intractable. We do not fully understand the reasons for this phenomenon, though the notions of NP completeness we will see later do offer a partial explanation. It may be related to the fact that optimizing a polynomial often amounts to equations on its derivative. The derivative of a quadratic polynomial is linear, while the derivative of a cubic is quadratic, and, as we will see, the difference between solving linear and quadratic equations can be quite profound.
11.2.2 Solving linear equations

One of the most useful problems that people have been solving time and again is solving \( n \) linear equations in \( n \) variables. That is, solve equations of the form

\[
\begin{align*}
    a_{0,0} x_0 + a_{0,1} x_1 + \cdots + a_{0,n-1} x_{n-1} &= b_0 \\
    a_{1,0} x_0 + a_{1,1} x_1 + \cdots + a_{1,n-1} x_{n-1} &= b_1 \\
    \vdots &+ \vdots + \vdots & = \vdots \\
    a_{n-1,0} x_0 + a_{n-1,1} x_1 + \cdots + a_{n-1,n-1} x_{n-1} &= b_{n-1}
\end{align*}
\]

(11.4)

where \( \{a_{i,j}\}_{i,j \in [n]} \) and \( \{b_i\}_{i \in [n]} \) are real (or rational) numbers. More compactly, we can write this as the equations \( A x = b \) where \( A \) is an \( n \times n \) matrix, and we think of \( x, b \) are column vectors in \( \mathbb{R}^n \).

The standard Gaussian elimination algorithm can be used to solve such equations in polynomial time (i.e., determine if they have a solution, and if so, to find it). As we discussed above, if we are willing to allow some loss in precision, we even have algorithms that handle linear inequalities, also known as linear programming. In contrast, if we insist on integer solutions, the task of solving for linear equalities or inequalities is known as integer programming, and the best known algorithms are exponential time in the worst case.

**Remark 11.4** — Bit complexity of numbers. Whenever we discuss problems whose inputs correspond to numbers, the input length corresponds to how many bits are needed to describe the number (or, as is equivalent up to a constant factor, the number of digits in base 10, 16 or any other constant). The difference between the length of the input and the magnitude of the number itself can be of course quite profound. For example, most people would agree that there is a huge difference between having a billion (i.e. \( 10^9 \) dollars and having nine dollars. Similarly there is a huge difference between an algorithm that takes \( n \) steps on an \( n \)-bit number and an algorithm that takes \( 2^n \) steps.

One example, is the problem (discussed below) of finding the prime factors of a given integer \( N \). The natural algorithm is to search for such a factor by trying all numbers from 1 to \( N \), but that would take \( N \) steps which is exponential in the input length, which is number of bits needed to describe \( N \). (The running time of this algorithm can be easily improved to roughly \( \sqrt{N} \), but this is still exponential (i.e., \( 2^{n/2} \)) in the number \( n \) of bits to describe \( N \).) It is an important and long open question whether there is such an algorithm that runs in time polynomial in the input length (i.e., polynomial in \( \log N \)).
11.2.3 Solving quadratic equations
Suppose that we want to solve not just linear but also equations involving quadratic terms of the form $a_{i,j,k}x_j x_k$. That is, suppose that we are given a set of quadratic polynomials $p_1, \ldots, p_m$ and consider the equations $\{p_i(x) = 0\}$. To avoid issues with bit representations, we will always assume that the equations contain the constraints $\{x_i^2 \equiv x_i = 0\}_{i \in [n]}$. Since only 0 and 1 satisfy the equation $a^2 \equiv a$, this assumption means that we can restrict attention to solutions in $\{0, 1\}^n$.

Solving quadratic equations in several variable is a classical and extremely well motivated problem. This is the generalization of the classical case of single-variable quadratic equations that generations of high school students grapple with. It also generalizes the quadratic assignment problem, introduced in the 1950's as a way to optimize assignment of economic activities. Once again, we do not know a much better algorithm for this problem than the one that enumerates over all the $2^n$ possibilities.

11.3 MORE ADVANCED EXAMPLES
We now list a few more examples of interesting problems that are a little more advanced but are of significant interest in areas such as physics, economics, number theory, and cryptography.

11.3.1 Determinant of a matrix
The determinant of a $n \times n$ matrix $A$, denoted by $\det(A)$, is an extremely important quantity in linear algebra. For example, it is known that $\det(A) \neq 0$ if and only if $A$ is nonsingular, which means that it has an inverse $A^{-1}$, and hence we can always uniquely solve equations of the form $Ax = b$ where $x$ and $b$ are $n$-dimensional vectors. More generally, the determinant can be thought of as a quantitative measure as to what extent $A$ is far from being singular. If the rows of $A$ are “almost” linearly dependent (for example, if the third row is very close to being a linear combination of the first two rows) then the determinant will be small, while if they are far from it (for example, if they are are orthogonal to one another, then the determinant will be large). In particular, for every matrix $A$, the absolute value of the determinant of $A$ is at most the product of the norms (i.e., square root of sum of squares of entries) of the rows, with equality if and only if the rows are orthogonal to one another.

The determinant can be defined in several ways. One way to define the determinant of an $n \times n$ matrix $A$ is:

$$\det(A) = \sum_{\pi \in S_n} \text{sign}(\pi) \prod_{i \in [n]} A_{i, \pi(i)} \hspace{1cm} (11.5)$$
where $S_n$ is the set of all permutations from $[n]$ to $[n]$ and the sign of a permutation $\pi$ is equal to $-1$ raised to the power of the number of inversions in $\pi$ (pairs $i, j$ such that $i > j$ but $\pi(i) < \pi(j)$).

This definition suggests that computing $\det(A)$ might require summing over $|S_n|$ terms which would take exponential time since $|S_n| = n! > 2^n$. However, there are other ways to compute the determinant. For example, it is known that $\det$ is the only function that satisfies the following conditions:

1. $\det(AB) = \det(A)\det(B)$ for every square matrices $A, B$.

2. For every $n \times n$ triangular matrix $T$ with diagonal entries $d_0, \ldots, d_{n-1}$, $\det(T) = \prod_{i=0}^{n} d_i$. In particular $\det(I) = 1$ where $I$ is the identity matrix. (A triangular matrix is one in which either all entries below the diagonal, or all entries above the diagonal, are zero.)

3. $\det(S) = -1$ where $S$ is a “swap matrix” that corresponds to swapping two rows or two columns of $I$. That is, there are two coordinates $a, b$ such that for every $i, j$, $S_{i,j} = \begin{cases} 1 & i = j, i \notin \{a, b\} \\ 1 & \{i, j\} = \{a, b\} \\ 0 & \text{otherwise} \end{cases}$.

Using these rules and the Gaussian elimination algorithm, it is possible to tell whether $A$ is singular or not, and in the latter case, decompose $A$ as a product of a polynomial number of swap matrices and triangular matrices. (Indeed one can verify that the row operations in Gaussian elimination corresponds to either multiplying by a swap matrix or by a triangular matrix.) Hence we can compute the determinant for an $n \times n$ matrix using a polynomial time of arithmetic operations.

11.3.2 Permanent of a matrix

Given an $n \times n$ matrix $A$, the permanent of $A$ is defined as

$$ perm(A) = \sum_{\pi \in S_n} \prod_{i \in [n]} A_{i, \pi(i)} \quad (11.6) $$

That is, $perm(A)$ is defined analogously to the determinant in (11.5) except that we drop the term sign$(\pi)$. The permanent of a matrix is a natural quantity, and has been studied in several contexts including combinatorics and graph theory. It also arises in physics where it can be used to describe the quantum state of multiple Boson particles (see here and here).
Permanent modulo 2. If the entries of $A$ are integers, then we can define the Boolean function $\text{perm}_2$ which outputs on input a matrix $A$ the result of the permanent of $A$ modulo 2. It turns out that we can compute $\text{perm}_2(A)$ in polynomial time. The key is that modulo 2, $-x$ and $+x$ are the same quantity and hence, since the only difference between (11.5) and (11.6) is that some terms are multiplied by $-1$, \[ \det(A) \mod 2 = \text{perm}(A) \mod 2 \] for every $A$.

Permanent modulo 3. Emboldened by our good fortune above, we might hope to be able to compute the permanent modulo any prime $p$ and perhaps in full generality. Alas, we have no such luck. In a similar “two to three” type of a phenomenon, we do not know of a much better than brute force algorithm to even compute the permanent modulo 3.

11.3.3 Finding a zero-sum equilibrium

A zero sum game is a game between two players where the payoff for one is the same as the penalty for the other. That is, whatever the first player gains, the second player loses. As much as we want to avoid them, zero sum games do arise in life, and the one good thing about them is that at least we can compute the optimal strategy.

A zero sum game can be specified by an $n \times n$ matrix $A$, where if player 1 chooses action $i$ and player 2 chooses action $j$ then player one gets $A_{i,j}$ and player 2 loses the same amount. The famous Min Max Theorem by John von Neumann states that if we allow probabilistic or “mixed” strategies (where a player does not choose a single action but rather a distribution over actions) then it does not matter who plays first and the end result will be the same. Mathematically the min max theorem is that if we let $\Delta_n$ be the set of probability distributions over $[n]$ (i.e., non-negative columns vectors in $\mathbb{R}^n$ whose entries sum to 1) then

\[
\max_{p \in \Delta_n} \min_{q \in \Delta_n} p^\top A q = \min_{q \in \Delta_n} \max_{p \in \Delta_n} p^\top A q \quad (11.7)
\]

The min-max theorem turns out to be a corollary of linear programming duality, and indeed the value of (11.7) can be computed efficiently by a linear program.

11.3.4 Finding a Nash equilibrium

Fortunately, not all real-world games are zero sum, and we do have more general games, where the payoff of one player does not necessarily equal the loss of the other. John Nash won the Nobel prize for showing that there is a notion of equilibrium for such games as well. In many economic texts it is taken as an article of faith that when actual agents are involved in such a game then they reach a Nash
equilibrium. However, unlike zero sum games, we do not know of an efficient algorithm for finding a Nash equilibrium given the description of a general (non zero sum) game. In particular this means that, despite economists’ intuitions, there are games for which natural strategies will take exponential number of steps to converge to an equilibrium.

11.3.5 Primality testing
Another classical computational problem, that has been of interest since the ancient greeks, is to determine whether a given number \( N \) is prime or composite. Clearly we can do so by trying to divide it with all the numbers in \( 2, \ldots, N - 1 \), but this would take at least \( N \) steps which is exponential in its bit complexity \( n = \log N \). We can reduce these \( N \) steps to \( \sqrt{N} \) by observing that if \( N \) is a composite of the form \( N = PQ \) then either \( P \) or \( Q \) is smaller than \( \sqrt{N} \). But this is still quite terrible. If \( N \) is a 1024 bit integer, \( \sqrt{N} \) is about \( 2^{512} \), and so running this algorithm on such an input would take much more than the lifetime of the universe.

Luckily, it turns out we can do radically better. In the 1970’s, Rabin and Miller gave probabilistic algorithms to determine whether a given number \( N \) is prime or composite in time \( \text{poly}(n) \) for \( n = \log N \). We will discuss the probabilistic model of computation later in this course. In 2002, Agrawal, Kayal, and Saxena found a deterministic \( \text{poly}(n) \) time algorithm for this problem. This is surely a development that mathematicians from Archimedes till Gauss would have found exciting.

11.3.6 Integer factoring
Given that we can efficiently determine whether a number \( N \) is prime or composite, we could expect that in the latter case we could also efficiently find the factorization of \( N \). Alas, no such algorithm is known. In a surprising and exciting turn of events, the non existence of such an algorithm has been used as a basis for encryptions, and indeed it underlies much of the security of the world wide web. We will return to the factoring problem later in this course. We remark that we do know much better than brute force algorithms for this problem. While the brute force algorithms would require \( 2^{\Omega(n)} \) time to factor an \( n \)-bit integer, there are known algorithms running in time roughly \( 2^{O(\sqrt{n})} \) and also algorithms that are widely believed (though not fully rigorously analyzed) to run in time roughly \( 2^{O(n^{1/3})} \). (By “roughly” we mean that we neglect factors that are polylogarithmic in \( n \).)
11.4 OUR CURRENT KNOWLEDGE

The difference between an exponential and polynomial time algorithm might seem merely “quantitative” but it is in fact extremely significant. As we’ve already seen, the brute force exponential time algorithm runs out of steam very very fast, and as Edmonds says, in practice there might not be much difference between a problem where the best algorithm is exponential and a problem that is not solvable at all. Thus the efficient algorithms we mention above are widely used and power many computer science applications. Moreover, a polynomial-time algorithm often arises out of significant insight to the problem at hand, whether it is the “max-flow min-cut” result, the solvability of the determinant, or the group theoretic structure that enables primality testing. Such insight can be useful regardless of its computational implications.

At the moment we do not know whether the “hard” problems are truly hard, or whether it is merely because we haven’t yet found the right algorithms for them. However, we will now see that there are problems that do inherently require exponential time. We just don’t know if any of the examples above fall into that category.

Lecture Recap

- There are many natural problems that have polynomial-time algorithms, and other natural problems that we’d love to solve, but for which the best known algorithms are exponential.
- Often a polynomial time algorithm relies on discovering some hidden structure in the problem, or finding a surprising equivalent formulation for it.
- There are many interesting problems where there is an exponential gap between the best known algorithm and the best algorithm that we can rule out. Closing this gap is one of the main open questions of theoretical computer science.

11.5 EXERCISES

Exercise 11.1 — exponential time algorithm for longest path. The naive algorithm for computing the longest path in a given graph could take more than \( n! \) steps. Give a \( poly(n)2^n \) time algorithm for the longest path problem in \( n \) vertex graphs.

Exercise 11.2 — 2SAT algorithm. For every 2CNF \( \varphi \), define the graph \( G_{\varphi} \) on \( 2n \) vertices corresponding to the literals \( x_1, \ldots, x_n, \overline{x}_1, \ldots, \overline{x}_n \), such that there is an edge \( \overline{\ell}_i \overline{\ell}_j \) iff the constraint \( \overline{\ell}_i \lor \ell_j \) is in \( \varphi \). Prove that \( \varphi \) holds if and only if there is a path from \( s \) to \( t \) in \( G_{\varphi} \). Do this iteratively for \( S \)'s of growing sizes.

\(^2\) Hint: Use dynamic programming to compute for every \( s, t \in [n] \) and \( S \subseteq [n] \) the value \( P(s, t, S) \) which equals 1 if there is a simple path from \( s \) to \( t \) that uses exactly the vertices in \( S \).
is unsatisfiable if and only if there is some $i$ such that there is a path from $x_i$ to $\overline{x_i}$ and from $\overline{x_i}$ to $x_i$ in $G_\varphi$. Show how to use this to solve 2SAT in polynomial time.

### 11.6 BIBLIOGRAPHICAL NOTES

The classic undergraduate introduction to algorithms text is [Cor+09]. Two texts that are less “encyclopedic” are Kleinberg and Tardos [KT06], and Dasgupta, Papadimitriou and Vazirani [DPV08]. Jeff Erickson’s book is an excellent algorithms text that is freely available online.

The origins of the minimum cut problem date to the cold war. Specifically, Ford and Fulkerson discovered their max-flow/min-cut algorithm in 1955 as a way to find out the minimum amount of train tracks that would need to be blown up to disconnect Russia from the rest of Europe. See the survey [Sch05] for more.

### 11.7 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

3 TODO: add reference to best algorithm for longest path - probably the Bjorklund algorithm
“When the measure of the problem-size is reasonable and when the sizes assume values arbitrarily large, an asymptotic estimate of ... the order of difficulty of [an] algorithm ... is theoretically important. It cannot be rigged by making the algorithm artificially difficult for smaller sizes”, Jack Edmonds, “Paths, Trees, and Flowers”, 1963

“The computational complexity of a sequence is to be measured by how fast a multitape Turing machine can print out the terms of the sequence. This particular abstract model of a computing device is chosen because much of the work in this area is stimulated by the rapidly growing importance of computation through the use of digital computers, and all digital computers in a slightly idealized form belong to the class of multitape Turing machines.”, Juris Hartmanis and Richard Stearns, “On the computational complexity of algorithms”, 1963.

In Chapter 11 we saw examples of efficient algorithms, and made some claims about their running time, but did not give a mathematically precise definition for this concept. We do so in this chapter, using the models of Turing machines and RAM machines (or equivalently NAND-TM and NAND-RAM) we have seen before. The running time of an algorithm is not a fixed number since any non-trivial algorithm will take longer to run on longer inputs. Thus, what we want to measure is the dependence between the number of steps the algorithms takes and the length of the input. In particular we care about the distinction between algorithms that take at most polynomial time (i.e., $O(n^c)$ time for some constant $c$) and problems for which every algorithm requires at least exponential time (i.e., $\Omega(2^{n^c})$ for some $c$). As mentioned in Edmond’s quote above, the difference between these
In this chapter we formally define the notion of a function being computable in $T(n)$ time where $T$ is some function mapping the length of the input to a bound on the number of computation steps. We then do the following (see also Fig. 12.1):

- Define the class $P$ of Boolean functions that can be computed in polynomial time and its superset $\text{EXP}$ of functions that can be computed in exponential time.

- Show that the time to compute a function using a Turing Machine and using a RAM machine (or NAND-RAM program) is polynomially related which in particular means that the classes $P$ and $\text{EXP}$ can be equivalently defined using either Turing Machines or RAM machines / NAND-RAM programs.

- Give an efficient universal NAND-RAM program and use this to establish the time hierarchy theorem that in particular implies that $P \subsetneq \text{EXP}$.

- We relate the notions defined here to the non uniform models of Boolean circuits and NAND-CIRC programs defined in Chapter 3. We define $P_{\text{poly}}$ to be the class of functions computed by a sequence of polynomial-sized circuits. We prove that $P \subseteq P_{\text{poly}}$ and that $P_{\text{poly}}$ contains uncomputable functions.

### 12.1 FORMALLY DEFINING RUNNING TIME

Our models of computation such Turing Machines, NAND-TM and NAND-RAM programs and others all operate by executing a sequence
of instructions on an input one step at a time. The running time is defined by measuring the number of steps as a function of the length of the input. We start by defining running time with respect to Turing Machines:

**Definition 12.1 — Running time (Turing Machines).** Let \( T : \mathbb{N} \to \mathbb{N} \) be some function mapping natural numbers to natural numbers. We say that a function \( F : \{0, 1\}^* \to \{0, 1\}^* \) is computable in \( T(n) \) Single-Tape-Turing-Machine time (TM-time for short) if there exists a Turing Machine \( M \) such that for every sufficiently large \( n \) and every \( x \in \{0, 1\}^n \), when given input \( x \), the machine \( M \) halts after executing at most \( T(n) \) steps and outputs \( F(x) \).

We define \( \text{TIME}_{TM}(T(n)) \) to be the set of Boolean functions (functions mapping \( \{0, 1\}^* \) to \( \{0, 1\} \)) that are computable in \( T(n) \) TM time.

**Solved Exercise 12.1 — Example of time bounds.** Prove that \( \text{TIME}_{TM}(10 \cdot n^3) \subseteq \text{TIME}_{TM}(2^n) \).

**Solution:**

The proof is illustrated in Fig. 12.2. Suppose that \( F \in \text{TIME}_{TM}(10\cdot n^3) \) and hence there some number \( N_0 \) and a machine \( M \) such that for every \( n > N_0 \), and \( x \in \{0, 1\}^* \), \( M(x) \) outputs \( F(x) \) within at

**Figure 12.2:** Comparing \( T(n) = 10n^3 \) with \( T'(n) = 2^n \) (on the right figure the Y axis is in log scale). Since for every large enough \( n \), \( T'(n) \geq T(n) \), \( \text{TIME}_{TM}(T(n)) \subseteq \text{TIME}_{TM}(T'(n)) \).
most $10 \cdot n^3$ steps. Since $10 \cdot n^3 = o(2^n)$, there is some number $N_1$ such that for every $n > N_1$, $10 \cdot n^3 < 2^n$. Hence for every $n > \max\{N_0, N_1\}$, $M(x)$ will output $F(x)$ within at most $2^n$ steps, just demonstrating that $F \in \text{TIME}_T(2^n)$.

12.1.1 Polynomial and Exponential Time

Unlike the notion of computability, the exact running time can be a function of the model we use. However, it turns out that if we only care about “coarse enough” resolution (as will most often be the case) then the choice of the model, whether Turing Machines, RAM machines, NAND-TM/NAND-RAM programs, or C/Python programs, does not matter. This is known as the extended Church-Turing Thesis. Specifically we will mostly care about the difference between polynomial and exponential time.

The two main time complexity classes we will be interested in are the following:

- **Polynomial time:** A function $F : \{0, 1\}^* \to \{0, 1\}$ is computable in polynomial time if it is in the class $P = \bigcup_{c \in \{1, 2, 3, \ldots\}} \text{TIME}(n^c)$. That is, $F \in P$ if there is an algorithm to compute $F$ that runs in time at most polynomial (i.e., at most $n^c$ for some constant $c$) in the length of the input.

- **Exponential time:** A function $F : \{0, 1\}^* \to \{0, 1\}$ is computable in exponential time if it is in the class $\text{EXP} = \bigcup_{c \in \{1, 2, 3, \ldots\}} \text{TIME}(2^{n^c})$. That is, $F \in \text{EXP}$ if there is an algorithm to compute $F$ that runs in time at most exponential (i.e., at most $2^{n^c}$ for some constant $c$) in the length of the input.

In other words, these are defined as follows:

**Definition 12.2 — $P$ and $\text{EXP}$**. Let $F : \{0, 1\}^* \to \{0, 1\}$. We say that $F \in P$ if there is a polynomial $p : \mathbb{N} \to \mathbb{R}$ and a Turing Machine $M$ such that for every $x \in \{0, 1\}^*$, $M(x)$ runs in at most $p(|x|)$ steps and outputs $F(x)$.

We say that $F \in \text{EXP}$ if there is a polynomial $p : \mathbb{N} \to \mathbb{R}$ and a Turing Machine $M$ such that for every $x \in \{0, 1\}^*$, $M(x)$ runs in at most $2^{p(|x|)}$ steps and outputs $F(x)$.

Please make sure you understand why Definition 12.2 and the bullets above define the same classes.

Sometimes students think of the class $\text{EXP}$ as corresponding to functions that are not in $P$. However, this is not the case. If $F$ is in $\text{EXP}$ then it can be computed.
Since exponential time is much larger than polynomial time, $P \subseteq \text{EXP}$. All of the problems we listed in Chapter 11 are in EXP, but as we’ve seen, for some of them there are much better algorithms that demonstrate that they are in fact in the smaller class $P$.

<table>
<thead>
<tr>
<th>$P$</th>
<th>EXP (but not known to be in $P$)</th>
</tr>
</thead>
<tbody>
<tr>
<td>Shortest path</td>
<td>Longest Path</td>
</tr>
<tr>
<td>Min cut</td>
<td>Max cut</td>
</tr>
<tr>
<td>2SAT</td>
<td>3SAT</td>
</tr>
<tr>
<td>Linear eqs</td>
<td>Quad. eqs</td>
</tr>
<tr>
<td>Zerosum</td>
<td>Nash</td>
</tr>
<tr>
<td>Determinant</td>
<td>Permanent</td>
</tr>
<tr>
<td>Primality</td>
<td>Factoring</td>
</tr>
</tbody>
</table>

Table: A table of the examples from Chapter 11. All these problems are in EXP but the only the ones on the left column are currently known to be in $P$ as well (i.e., they have a polynomial-time algorithm). See also Fig. 12.3.

**Remark 12.3 — Boolean versions of problems.** Many of the problems defined in Chapter 11 correspond to non-Boolean functions (functions with more than one bit of output) while $P$ and EXP are sets of Boolean functions. However, for every non-Boolean function $F$ we can always define a computationally-equivalent Boolean function $G$ by letting $G(x, i)$ be the $i$-th bit of $F(x)$ (see Exercise 12.3). Hence the table above, as well as Fig. 12.3, refer to the computationally-equivalent Boolean variants of these problems.

### 12.2 Modeling Running Time Using RAM Machines / NAND-RAM

Turing Machines are a clean theoretical model of computation, but do not closely correspond to real-world computing architectures. The discrepancy between Turing Machines and actual computers does not matter much when we consider the question of which functions are computable, but can make a difference in the context of efficiency. Even a basic staple of undergraduate algorithms such as “merge sort” cannot be implemented on a Turing Machine in $O(n \log n)$ time (see Section 12.8). RAM machines (or equivalently, NAND-RAM programs)
more closely match actual computing architecture and what we mean when we say $O(n)$ or $O(n \log n)$ algorithms in algorithms courses or whiteboard coding interviews. We can define running time with respect to NAND-RAM programs just as we did for Turing Machines.

**Definition 12.4 — Running time (RAM).** Let $T : \mathbb{N} \to \mathbb{N}$ be some function mapping natural numbers to natural numbers. We say that a function $F : \{0, 1\}^* \to \{0, 1\}^*$ is computable in $T(n)$ RAM time (RAM-time for short) if there exists a NAND-RAM program $P$ such that for every sufficiently large $n$ and every $x \in \{0, 1\}^n$, when given input $x$, the program $P$ halts after executing at most $T(n)$ iterations and outputs $F(x)$.

We define $\text{TIME}_{\text{RAM}}(T(n))$ to be the set of Boolean functions (functions mapping $\{0, 1\}^*$ to $\{0, 1\}$) that are computable in $T(n)$ TM time.

Because NAND-RAM programs correspond more closely to our natural notions of running time, we will use NAND-RAM as our “default” model of running time, and hence use $\text{TIME}(T(n))$ (without any subscript) to denote $\text{TIME}_{\text{RAM}}(T(n))$. However, it turns out that as long as we only care about the difference between exponential and polynomial time, this does not make much difference. The reason is that Turing Machines can simulate NAND-RAM programs with at most a polynomial overhead (see also Fig. 12.4):

**Theorem 12.5 — Relating RAM and Turing machines.** Let $T : \mathbb{N} \to \mathbb{N}$ be a function such that $T(n) \geq n$ for every $n$ and the map $n \mapsto T(n)$ can be computed by a Turing machine in time $O(T(n)^3)$. Then

$$\text{TIME}_{\text{TM}}(T(n)) \subseteq \text{TIME}_{\text{RAM}}(T(n)) \subseteq \text{TIME}_{\text{TM}}(T(n)^4).$$

(12.1)

Big Idea 12 While the precise definition of running time can depend on the computational model, as long as we only care about the distinction between polynomial and exponential, all the models we considered are equivalent to each other.

All non pathological time bound functions such as $T(n) = n$, $T(n)n \log n$, $T(n) = 2^n$ etc. satisfy the conditions of Theorem 12.5, see also Remark 12.6. The constant 4 can be improved to a smaller value, though this will not be important for us. Theorem 12.5 implies that the P and EXP could have been equivalently defined using NAND-RAM programs instead of Turing Machines, as they would have contained the exact same set of functions. Similar equivalence results are known for many models including cellular automata, C/Python/Javascript.

Figure 12.4: The proof of Theorem 12.5 shows that we can simulate $T$ steps of a Turing Machine with $T$ steps of a NAND-RAM program, and can simulate $T$ steps of a NAND-RAM program with $o(T^4)$ steps of a Turing Machine. Hence $\text{TIME}_{\text{TM}}(T(n)) \subseteq \text{TIME}_{\text{RAM}}(T(n)) \subseteq \text{TIME}_{\text{TM}}(T(n)^4)$. 
programs, parallel computers, and a great many other models, which justifies the choice of P as capturing a technology-independent notion of tractability. This equivalence between Turing machines and NAND-RAM (as well as other models) allows us to pick our favorite model depending on the task at hand (i.e., “have our cake and eat it too”). When we want to design an algorithm, we can use the extra power and convenience afforded by NAND-RAM. When we want to analyze a program or prove a negative result, we can restrict attention to Turing machines.

**Proof Idea:**

The direction $\text{TIME}_{\text{TM}}(T(n)) \subseteq \text{TIME}_{\text{RAM}}(T(n))$ is not hard to show, since a NAND-RAM program can simulate a Turing Machine so that each step of the machine is captured by one iteration of the main loop of the program. Thus the heart of the theorem is to prove that $\text{TIME}_{\text{RAM}}(T(n)) \subseteq \text{TIME}_{\text{TM}}(T(n)^4)$. This proof closely follows the proof of Theorem 7.1, where we have shown that every function $F$ that is computable by a NAND-RAM program $P$ is computable by a Turing Machine (or equivalently a NAND-TM program) $M$. To prove Theorem 12.5, we follow the exact same proof but just check that the overhead of the simulation of $P$ by $M$ is polynomial. The proof has many details, but is not deep. It is therefore much more important that you understand the statement of this theorem than its proof.

**Proof of Theorem 12.5.** We only focus on the non-trivial direction $\text{TIME}_{\text{RAM}}(T(n)) \subseteq \text{TIME}_{\text{TM}}(T(n)^4)$. Let $F \in \text{TIME}_{\text{RAM}}(T(n))$. $F$ can be computed in time $T(n)$ by some NAND-RAM program $P$ and we need to show that it can also be computed in time $T(n)^4$ by a Turing Machine $M$. This will follow from showing that $F$ can be computed in time $T(n)^4$ by a NAND-TM program, since for every NAND-TM program $Q$ there is a Turing Machine $M$ simulating it such that each iteration of $Q$ corresponds to a single step of $M$.

As mentioned above, we follow the proof of Theorem 7.1 (simulation of NAND-RAM programs using NAND-TM programs) and use the exact same simulation, but with a more careful accounting of the number of steps that the simulation costs. Recall, that the simulation of NAND-RAM works by “peeling off” features of NAND-RAM one by one, until we are left with NAND-TM.

We will not provide the full details but will present the main ideas used in showing that every feature of NAND-RAM can be simulated by NAND-TM with at most a polynomial overhead:

1. Recall that every NAND-RAM variable or array element can contain an integer between 0 and $T$ where $T$ is the number of lines that
have been executed so far. Therefore if $P$ is a NAND-RAM program that computes $F$ in $T(n)$ time, then on inputs of length $n$, all integers used by $P$ are of magnitude at most $T(n)$. This means that the largest value $i$ can ever reach is at most $T(n)$ and so each one of $P$’s variables can be thought of as an array of at most $T(n)$ indices, each of which holds a natural number of magnitude at most $T(n)$. We let $\ell = \lceil \log T(n) \rceil$ be the number of bits needed to encode such numbers. (We can start off the simulation by computing $T(n)$ and $\ell$.)

2. We can encode a NAND-RAM array of length $\leq T(n)$ containing numbers in $\{0, \ldots, T(n) - 1\}$ as a Boolean (i.e., NAND-TM) array of $T(n)\ell = O(T(n) \log T(n))$ bits, which we can also think of as a two dimensional array as we did in the proof of Theorem 7.1. We encode a NAND-RAM scalar containing a number in $\{0, \ldots, T(n) - 1\}$ simply by a shorter NAND-TM array of $\ell$ bits.

3. We can simulate the two dimensional arrays using one-dimensional arrays of length $T(n)\ell = O(T(n) \log T(n))$. All the arithmetic operations on integers use the grade-school algorithms, that take time that is polynomial in the number $\ell$ of bits of the integers, which is poly($\log T(n)$) in our case. Hence we can simulate $T(n)$ steps of NAND-RAM with $O(T(n)\text{poly}(\log T(n))$ steps of a model that uses random access memory but only Boolean-valued one-dimensional arrays.

4. The most expensive step is to translate from random access memory to the sequential memory model of NAND-TM/Turing Machines. As we did in the proof of Theorem 7.1 (see Section 7.2), we can simulate accessing an array $\text{Foo}$ at some location encoded in an array $\text{Bar}$ by:

a. Copying $\text{Bar}$ to some temporary array $\text{Temp}$

b. Having an array $\text{Index}$ which is initially all zeros except 1 at the first location.

c. Repeating the following until $\text{Temp}$ encodes the number 0:
   - Decrease the number encoded temp by 1. (*Take number of steps polynomial in $\ell = \lceil \log T(n) \rceil$.*)
   - Decrease $i$ until it is equal to 0. (*Take $O(T(n))$ steps.*)
   - Scan $\text{Index}$ until we reach the point in which it equals 1 and then change this 1 to 0 and go one step further and write 1 in this location. (*Takes $O(T(n))$ steps.*)

d. When we are done we know that if we scan $\text{Index}$ until we reach the point in which $\text{Index}[i] = 1$ then $i$ contains the value that was encoded by $\text{Bar}$ (*Takes $O(T(n))$ steps.*)
The total cost for each such operation is \( O(T(n)^2 + T(n)\text{poly}(\log T(n))) = O(T(n)^2) \) steps.

In sum, we simulate a single step of NAND-RAM using \( O(T(n)^2\text{poly}(\log T(n))) \) steps of NAND-TM, and hence the total simulation time is \( O(T(n)^3\text{poly}(\log T(n))) \) which is smaller than \( T(n)^4 \) for sufficiently large \( n \).

**Remark 12.6 — Nice time bounds.** When considering general time bounds such we need to make sure to rule out some “pathological” cases such as functions \( T \) that don’t give enough time for the algorithm to read the input, or functions where the time bound itself is uncomputable. We say that a function \( T : \mathbb{N} \to \mathbb{N} \) is a nice time bound function (or nice function for short)

if for every \( n \in \mathbb{N}, T(n) \geq n \) (i.e., \( T \) allows enough time to read the input), for every \( n' \geq n, T(n') \geq T(n) \) (i.e., \( T \) allows more time on longer inputs), and the map \( F(x) = 1^{T(|x|)} \) (i.e., mapping a string of length \( n \) to a sequence of \( T(n) \) ones) can be computed by a NAND-RAM program in \( O(T(n)) \) time.

All the “normal” time complexity bounds we encounter in applications such as \( T(n) = 100n, T(n) = n^2 \log n, T(n) = 2^n \pi \), etc. are “nice”.

Hence from now on we will only care about the class \( \text{TIME}(T(n)) \) when \( T \) is a “nice” function. The computability condition is in particular typically easily satisfied. For example, for arithmetic functions such as \( T(n) = n^3 \), we can typically compute the binary representation of \( T(n) \) in time polynomial in the number of bits of \( T(n) \) and hence poly-logarithmic in \( T(n) \). Hence the time to write the string \( 1^{T(n)} \) in such cases will be \( T(n) + \text{poly}(\log T(n)) = O(T(n)) \).

### 12.3 Extended Church-Turing Thesis (Discussion)

Theorem 12.5 shows that the computational models of *Turing Machines* and *RAM Machines* / *NAND-RAM programs* are equivalent up to polynomial factors in the running time. Other examples of polynomially equivalent models include:

- All standard programming languages, including C/Python/-JavaScript/Lisp/etc.
- The \( \lambda \) calculus (see also Section 12.8).
- Cellular automata
- Parallel computers
• Biological computing devices such as DNA-based computers.

The Extended Church Turing Thesis is the statement this is true for all physically realizable computing models. In other words, the extended Church Turing thesis says that for every scalable computing device $C$ (which has a finite description but can be in principle used to run computation on arbitrarily large inputs), there is some constant $a$ such that for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ that $C$ can compute on $n$ length inputs using an $S(n)$ amount of physical resources, $F$ is in $\text{TIME}(S(n)^a)$. This is a strengthening of the ("plain") Church-Turing Thesis, discussed in Section 7.8, which states that the set of computable functions is the same for all physically realizable models, but without requiring the overhead in the simulation between different models to be at most polynomial.

All the current constructions of scalable computational models and programming language conform to the Extended Church-Turing Thesis, in the sense that they can be with polynomial overhead by Turing Machines (and hence also by NAND-TM or NAND-RAM programs). consequently, the classes $P$ and $\text{EXP}$ are robust to the choice of model, and we can use the programming language of our choice, or high level descriptions of an algorithm, to determine whether or not a problem is in $P$.

Like the Church-Turing thesis itself, the extended Church-Turing thesis is in the asymptotic setting and does not directly yield an experimentally testable prediction. However, it can be instantiated with more concrete bounds on the overhead, yielding experimentally-testable predictions such as the Physical Extended Church-Turing Thesis we mentioned in Section 5.6.

In the last hundred+ years of studying and mechanizing computation, no one has yet constructed a scalable computing device (or even gave a convincing blueprint) that violates the extended Church Turing Thesis. However, quantum computing, if realized, will pose a serious challenge to the extended Church-Turing Thesis (see Chapter 22). However, even if the promises of quantum computing are fully realized, the extended Church-Turing thesis is “morally” correct, in the sense that, while we do need to adapt the thesis to account for the possibility of quantum computing, its broad outline remains unchanged. We are still able to model computation mathematically, we can still treat programs as strings and have a universal program, we still have hierarchy and uncomputability results, and there is still no reason to doubt the ("plain") Church-Turing thesis. Moreover, the prospect of quantum computing does not seem to make a difference for the time complexity of many (though not all!) of the concrete problems that we care about. In particular, out of all the example problems men-
tioned in Chapter 11, as far as we know, the complexity of only one—integer factoring—is affected by modifying our model to include quantum computers as well.

12.4 EFFICIENT UNIVERSAL MACHINE: A NAND-RAM INTERPRETER IN NAND-RAM

We have seen in Theorem 8.1 the “universal Turing Machine”. Examining that proof, and combining it with Theorem 12.5, we can see that the program $U$ has a polynomial overhead, in the sense that it can simulate $T$ steps of a given NAND-TM (or NAND-RAM) program $P$ on an input $x$ in $O(T^4)$ steps. But in fact, by directly simulating NAND-RAM programs we can do better with only a constant multiplicative overhead. That is, there is a universal NAND-RAM program $U$ such that for every NAND-RAM program $P$, $U$ simulates $T$ steps of $P$ using only $O(T)$ steps.

Theorem 12.7 — Efficient universality of NAND-RAM. There exists a NAND-RAM program $U$ satisfying the following:

1. ($U$ is a universal NAND-RAM program.) For every NAND-RAM program $P$ and input $x$, $U(P, x) = P(x)$ where by $U(P, x)$ we denote the output of $U$ on a string encoding the pair $(P, x)$.

2. ($U$ is efficient.) For every NAND-RAM program $P$ there is some constant $C$ such that if $P$ halts on input $x$ after most $T$ iterations of its loop, then $U(P, x)$ halts after at most $C \cdot T$ iterations of its loop.

Proof of Theorem 12.7. To present a universal NAND-RAM program in full we would need to describe a precise representation scheme, as well as the full NAND-RAM instructions for the program. While this can be done, it is more important to focus on the main ideas, and

Before reading the proof of Theorem 12.7, try to think how you would try to write such an interpreter for NAND-RAM in your favorite programming language. You will likely find that your program requires $O(T)$ steps to perform this simulation. As in the case of Theorem 12.5, the proof of Theorem 12.7 is not very deep and it more important to understand its statement. If you understand how you would go about writing an interpreter for NAND-RAM using a modern programming language such as Python, then you know everything you need to know about this theorem.
so we just sketch the proof here. A specification of NAND-RAM is given in the Appendix, and for the purposes of this simulation, we can simply use the representation of the code NAND-RAM as an ASCII string.

The program \( U \) gets as input a NAND-RAM program \( P \) and an input \( x \) and simulates \( P \) one step at a time. To do so, \( U \) will do the following:

1. \( U \) will maintain variables `current_line`, and `number_steps` for the current line to be executed and the number of steps executed so far.

2. \( U \) will scan the code of \( P \) to find the number \( t \) of unique variable names that \( P \) uses. If we denote these names by \( \text{var}_0, \ldots, \text{var}_{t-1} \) then \( U \) maintains an array `Var_numbers` that contains a list of pairs of the form \((\text{var}_s, s)\) for \( s \in [t] \). Using `Var_numbers` we can translate the name of a variable to a number in \([t]\) that corresponds to its index.

3. \( U \) will maintain an array `Var_values` that will contain the current values of all \( P \)'s variables. If the \( s \)-th variable of \( P \) is a scalar variable, then its value will be stored in location `Var_values[s]`. If it is an array variable then the value of its \( i \)-th element will be stored in location `Var_values[t \cdot i + s]`.

4. To simulate a single step of \( P \), the program \( U \) will recover the line corresponding to `line_counter` and execute it. Since NAND-RAM has a constant number of arithmetic operations, we can simulate choosing which operation to execute with a sequence of a constantly many if-then-else's. When executing these operations, \( U \) will use the variable `step_counter` that keeps track of the iteration counter of \( P \).

Simulating a single step of \( P \) will take \( O(|P|) \) steps for the program \( U \) where \( |P| \) is the length of the description of \( P \) as a string (which in particular is an upper bound on the number \( t \) of variable \( P \) uses). Hence the total running time will be \( O(|P|T) \) which is \( O(T) \) when suppressing constants that depend on the program \( P \).

To be a little more concrete, here is some “pseudocode” description of the program \( U \):

```python
def U(P, x):
    t = number_variable_identifiers(P)  # number of distinct identifiers used in P
    # ...
    L = number_lines(P)
```
# denote names of P’s variables as var_0,...,
	var_(t-1)
Var.numbers = array encoding list [  
	(var_0,0),...,(var_(t-1),t-1)]
# Var.numbers: encode variable identifiers as number  
	0...t-1

Var.values = unbounded array initialized to 0
# if s in [t] corresponds to scalar then Var.values[s]  
is value of variable corresponding to s.
# if s corresponds to array then Var.values[t*i+s] is  
value of variable corresponding to s at position i

def varid(name):
    # scan the array Var.numbers and  
    # return the number between 0 and t-1
...

def get_scalar_value(name):
    return Var.values[varid(name)]

def get_array_value(name,i):
    return Var.values[t*i+varid(name)]

def set_scalar_value(name,val):
    Var.values[varid(name)] = val

def set_array_value(name,i,val):
    Var.values[t*i+varid(name)] = val

for i=0..|x|-1:
    set_array_value("X",i,x[i])
    set_array_value("X_nonblank",i,1)

current_line = 0
number_steps = 0

do {
    line = P[current_line] # extract current line of P
    # code to execute line
    # We use get/set procedures above to update vars
    ...  
    # Update counters
current_line = current_line + 1 (mod L)
number_steps = number_steps + 1
}
}

} until get_scalar_value("loop") == 0

# Produce output:
if get_scalar_value("loop") == 1: return "FAIL"
m = smallest m s.t. get_array_value("Y_nonblank",m) = 0
return [get_array_value("Y",i) for i=0..m-1]

12.5 THE TIME HIERARCHY THEOREM

Some functions are uncomputable, but are there functions that can be computed, but only at an exorbitant cost? For example, is there a function that can be computed in time $2^n$, but cannot be computed in time $2^{0.9n}$? It turns out that the answer is Yes:

**Theorem 12.8 — Time Hierarchy Theorem.** For every nice function $T$, there is a function $F : \{0,1\}^* \rightarrow \{0,1\}$ in $\text{TIME}(T(n) \log n) \setminus \text{TIME}(T(n))$.

There is nothing special about $\log n$, and we could have used any other efficiently computable function that tends to infinity with $n$.

![Figure 12.5: The Time Hierarchy Theorem (Theorem 12.8) states that all of these classes are distinct.](image)

**Proof Idea:**

In the proof of Theorem 8.4 (the uncomputability of the Halting problem), we have shown that the function $\text{HALT}$ cannot be computed in any finite time. An examination of the proof shows that it gives something stronger. Namely, the proof shows that if we fix our
computational budget to be $T$ steps, then not only we can’t distinguish between programs that halt and those that do not, but cannot even distinguish between programs that halt within at most $T'$ steps and those that take more than that (where $T'$ is some number depending on $T$). Therefore, the proof of Theorem 12.8 follows the ideas of the uncomputability of the halting problem, but again with a more careful accounting of the running time.

\[ \star \]

**Proof of Theorem 12.8.** Recall the Halting function $HALT : \{0, 1\}^* \rightarrow \{0, 1\}$ that was defined as follows: $HALT(P, x)$ equals 1 for every program $P$ and input $x$ s.t. $P$ halts on input $x$, and is equal to 0 otherwise. We cannot use the Halting function of course, as it is uncomputable and hence not in $TIME(T'(n))$ for any function $T'$. However, we will use the following variant of it:

We define the Bounded Halting function $HALT_T(P, x)$ to equal 1 for every NAND-RAM program $P$ such that $|P| \leq \log\log |x|$, and such that $P$ halts on the input $x$ within $100T(|x|)$ steps. $HALT_T$ equals 0 on all other inputs.\(^1\)

Theorem 12.8 is an immediate consequence of the following two claims:

**Claim 1:** $HALT_T \in TIME(T(n) \log n)$

and

**Claim 2:** $HALT_T \notin TIME(T(n))$.

Please make sure you understand why indeed the theorem follows directly from the combination of these two claims. We now turn to proving them.

**Proof of claim 1:** We can easily check in linear time whether an input has the form $P, x$ where $|P| \leq \log\log |x|$. Since $T(\cdot)$ is a nice function, we can evaluate it in $O(T(n))$ time. Thus, we can perform the check above, compute $T(|P| + |x|)$ and use the universal NAND-RAM program of Theorem 12.7 to evaluate $HALT_T$ in at most $poly(|P|)T(n)$ steps.\(^2\) Since $(\log\log n)^a = o(\log n)$ for every $a$, this will be smaller than $T(n)\log n$ for every sufficiently large $n$, hence completing the proof.

**Proof of claim 2:** This proof is the heart of Theorem 12.8, and is very reminiscent of the proof that $HALT$ is not computable. Assume, toward the sake of contradiction, that there is some NAND-RAM program $P^*$ that computes $HALT_T(P, x)$ within $T(|P| + |x|)$ steps. We are going to show a contradiction by creating a program $Q$ and showing that under our assumptions, if $Q$ runs for less than $T(n)$ steps when given (a padded version of) its own code as input then it actually runs for more than $T(n)$ steps and vice versa. (It is worth re-reading the last sentence twice or thrice to make sure you understand

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\(^{1}\) The constant 100 and the function $\log\log n$ are rather arbitrary, and are chosen for convenience in this proof.

\(^{2}\) Recall that we use $poly(m)$ to denote a quantity that is bounded by $am^b$ for some constants $a$, $b$ and every sufficiently large $m$. 
this logic. It is very similar to the direct proof of the uncomputability of the halting problem where we obtained a contradiction by using an assumed “halting solver” to construct a program that, given its own code as input, halts if and only if it does not halt.

We will define $Q$ to be the program that on input a string $z$ does the following:

1. If $z$ does not have the form $z = P1^m$ where $P$ represents a NAND-RAM program and $|P| < 0.1 \log \log m$ then return 0. (Recall that $1^m$ denotes the string of $m$ ones.)
2. Compute $b = P^*(P, z)$ (at a cost of at most $T(|P| + |z|)$ steps, under our assumptions).
3. If $b = 1$ then $Q$ goes into an infinite loop, otherwise it halts.

We chose $m$ sufficiently large so that $|Q| < 0.001 \log \log m$ where $|Q|$ denotes the length of the description of $Q$ as a string. We will reach a contradiction by splitting into cases according to whether or not $\text{HALT}_T(Q, Q1^m) = 0$ or 1.

On the one hand, if $\text{HALT}_T(Q, Q1^m) = 1$, then under our assumption that $P^*$ computes $\text{HALT}_T$, $Q$ will go into an infinite loop on input $z = Q1^m$, and hence in particular $Q$ does not halt within $100T(|Q| + m)$ steps on the input $z$. But this contradicts our assumption that $\text{HALT}_T(Q, Q1^m) = 1$.

This means that it must hold that $\text{HALT}_T(Q, Q1^m) = 0$. But in this case, since we assume $P^*$ computes $\text{HALT}_T$, $Q$ does not do anything in phase 3 of its computation, and so the only computation costs come in phases 1 and 2 of the computation. It is not hard to verify that Phase 1 can be done in linear and in fact less than $5|z|$ steps. Phase 2 involves executing $P^*$, which under our assumption requires $T(|Q| + m)$ steps. In total we can perform both phases in less than $100T(|Q| + m)$ in steps, which by definition means that $\text{HALT}_T(Q, Q1^m) = 1$, but this is of course a contradiction. This completes the proof of Claim 2 and hence of Theorem 12.8.

Solved Exercise 12.2 — P vs EXP. Prove that $P \subsetneq \text{EXP}$.  

Solution:

We need to show that there exists $F \in \text{EXP} \setminus P$. Let $T(n) = n^{\log n}$ and $T'(n) = n^{\log n/2}$. Both are nice functions. Since $T(n) / T'(n) = \omega(\log n)$, by the time hierarchy theorem, there exists some $F$ in $\text{TIME}(T'(n)) \subsetneq \text{TIME}(T(n))$. Since for sufficiently large $n$, $2^n > n^{\log n}$, $F \in \text{TIME}(2^n) \subset \text{EXP}$. On the other hand, $F \notin P$.  


Indeed, suppose otherwise that there was a constant $c > 0$ and a Turing Machine computing $F$ on $n$-length input in at most $n^c$ steps for all sufficiently large $n$. Then since for $n$ large enough $n^c < n^\log n/2$, it would have followed that $F \in \text{TIME}(n^\log n/2)$ contradicting our choice of $F$.

The time hierarchy theorem tells us that there are functions we can compute in $O(n^2)$ time but not $O(n)$, in $2^n$ time, but not $2^{\sqrt{n}}$, etc. In particular there are most definitely functions that we can compute in time $2^n$ but not $O(n)$. We have seen that we have no shortage of natural functions for which the best known algorithm requires roughly $2^n$ time, and that many people have invested significant effort in trying to improve that. However, unlike in the finite vs. infinite case, for all of the examples above at the moment we do not know how to rule out even an $O(n)$ time algorithm. We will however see that there is a single unproven conjecture that would imply such a result for most of these problems.

The time hierarchy theorem relies on the existence of an efficient universal NAND-RAM program, as proven in Theorem 12.7. For other models such as Turing Machines we have similar time hierarchy results showing that there are functions computable in time $T(n)$ and not in time $T(n)/f(n)$ where $f(n)$ corresponds to the overhead in the corresponding universal machine.

12.6 NON UNIFORM COMPUTATION

We have now seen two measures of "computation cost" for functions. In Section 4.6 we defined the complexity of computing finite functions using circuits / straightline programs. Specifically, for a finite function $g : \{0, 1\}^n \rightarrow \{0, 1\}$ and number $T \in \mathbb{N}, g \in \text{SIZE}(T)$ if there is circuit of at most $T$ NAND gates (or equivalently a $T$-line NAND-CIRC program) that computes $g$. To relate this to the classes $\text{TIME}(T(n))$ defined in this chapter we first need to extend the class $\text{SIZE}(T(n))$ from finite functions to functions with unbounded input length.

**Definition 12.9 — Non uniform computation.** Let $F : \{0, 1\}^* \rightarrow \{0, 1\}$ and $T : \mathbb{N} \rightarrow \mathbb{N}$ be a nice time bound. For every $n \in \mathbb{N}$, define $F|_n : \{0, 1\}^n \rightarrow \{0, 1\}$ to be the restriction of $F$ to inputs of size $n$. That is, $F|_n$ is the function mapping $\{0, 1\}^n$ to $\{0, 1\}$ such that for every $x \in \{0, 1\}^n$, $F|_n(x) = F(x)$.

We say that $F$ is non-uniformly computable in at most $T(n)$ size, denoted by $F \in \text{SIZE}(T(n))$ if there exists a sequence $(C_0, C_1, C_2, \ldots)$ of NAND circuits such that:
• For every $n \in \mathbb{N}$, $C_n$ computes the function $F_{|n}$
• For every sufficiently large $n$, $C_n$ has at most $T(n)$ gates.

The non uniform analog to the class $P$ is the class $P/poly$ defined as

$$P/poly = \bigcup_{c \in \mathbb{N}} SIZE(n^c).$$

There is a big difference between non uniform computation and uniform complexity classes such as $TIME(T(n))$ or $P$. The condition $F \in P$ means that there is a single Turing machine $M$ that computes $F$ on all inputs in polynomial time. The condition $F \in P/poly$ only means that for every input length $n$ there can be a different circuit $C_n$ that computes $F$ using polynomially many gates on inputs of these lengths. As we will see, $F \in P/poly$ does not necessarily imply that $F \in P$. However, the other direction is true:

**Theorem 12.10 — Nonuniform computation contains uniform computation.** There is some $a \in \mathbb{N}$ s.t. for every nice $T : \mathbb{N} \to \mathbb{N}$ and $F : \{0, 1\}^* \to \{0, 1\}$,

$$TIME(T(n)) \subseteq SIZE(T(n)^a).$$

In particular, Theorem 12.10 shows that for every $c$, $TIME(n^c) \subseteq SIZE(n^{ca})$ and hence $P \subseteq P/poly$.

**Proof Idea:**
The idea behind the proof is to “unroll the loop”. Specifically, we will use the programming language variants of non-uniform and uniform computation: namely NAND-CIRC and NAND-TM. The main difference between the two is that NAND-TM has loops. However, for every fixed $n$, if we know that a NAND-TM program runs in at most $T(n)$ steps, then we can replace its loop by simply “copying and pasting” its code $T(n)$ times, similar to how in Python we can replace code such as

```python
for i in range(4):
    print(i)
```

with the “loop free” code

```python
print(0)
print(1)
print(2)
print(3)
```

To make this idea into an actual proof we need to tackle one technical difficulty, and this is to ensure that the NAND-TM program is

![Figure 12.7: We can think of an infinite function $F : \{0, 1\}^* \to \{0, 1\}$ as a collection of finite functions $F_0, F_1, F_2, \ldots$ where $F_n : \{0, 1\}^n \to \{0, 1\}$ is the restriction of $F$ to inputs of length $n$. We say $F$ is in $P/poly$ if for every $n$, the function $F_{|n}$ is computable by a polynomial size NAND-CIRC program, or equivalently, a polynomial sized Boolean circuit.](image)
**oblivious** in the sense that the value of the index variable i in the j-th iteration of the loop will depend only on j and not on the contents of the input. We make a digression to do just that in Section 12.6.1 and then complete the proof of Theorem 12.10.

*  

### 12.6.1 Oblivious NAND-TM programs

Our idea to prove Theorem 12.10 involves “unrolling the loop”. That is, suppose that $P$ is a NAND-TM program of $k$ lines of the form

```plaintext
line_1
...
line_(k-1)
MODANDJMP(dir1,dir2)
```

and such that on every input $x \in \{0, 1\}^n$, $P$ runs for at most $T(n)$ iterations and returns the value $F(x)$.

We would want to transform it into a NAND-CIRC program for computing the finite function $F_l_n$ by simply dropping the last line “copying and pasting” $T(n)$ copies of the first $k - 1$ lines. However, we still need to decide what to do with arrays. Since the index variable i can move at most one step per iteration, it will never reach more than $T(n) - 1$ on inputs of length $n$. Hence we can replace an array Foo with $T = T(n)$ scalar variables foo_0, ..., foo_(T - 1). Now we would want to replace references to Foo[i] in the original NAND-TM program with references to a variable of the form foo_k for some number $k$ in the new NAND-CIRC program. We could do that if the original NAND-TM program $P$ had the property that in its j-th iteration, the value of the index variable i is always equal to the same number $k$, regardless of what was the input. This would mean that when we obtain a NAND-CIRC program by taking $T(n) - 1$ copies of $P$, we can replace all references of the form Foo[i] in the j-th copy with references to foo_k. A NAND-TM program with this property is called oblivious and we now show that it is possible to transform every NAND-TM program into one that is oblivious.

---

**Theorem 12.11 — Making NAND-TM oblivious.** Let $T : \mathbb{N} \to \mathbb{N}$ be a nice function and let $F \in \text{TIME}_{\text{TM}}(T(n))$. Then there is a NAND-TM program $P$ that computes $F$ in $O(T(n)^2)$ steps and satisfying the following. For every $n \in \mathbb{N}$ there is a sequence $i_0, i_1, ..., i_{m-1}$ such that for every $x \in \{0, 1\}^n$, if $P$ is executed on input $x$ then in the j-th iteration the variable i is equal to $i_j$.

In other words, Theorem 12.11 implies that if we can compute $F$ in $T(n)$ steps, then we can compute it in $O(T(n)^2)$ steps with a program
Figur 12.8: We simulate a $T(n)$-time NAND-TM program $P'$ with an oblivious NAND-TM program $P$ by adding special arrays $At\text{start}$ and $At\text{end}$ to mark positions 0 and $T(n) - 1$ respectively. The program $P$ will simply “sweep” its arrays from right to left and back again. If the original program $P'$ would have moved $i$ in a different direction then we wait $O(T)$ steps until we reach the same point back again, and so $P$ runs in $O(T \cdot n)$ time.

Proof Idea:
We can translate any NAND-TM program $P'$ into an oblivious program $P$ by making $P$ “sweep” its arrays. That is, the index $i$ in $P$ will always move all the way from position 0 to position $T(n) - 1$ and back again. We can then simulate the program $P'$ with at most $T(n)$ overhead: if $P'$ wants to move $i$ left when we are in a rightward sweep then we simply wait the at most $2T(n)$ steps until the next time we are back in the same position while sweeping to the left.

Proof of Theorem 12.11. Let $P'$ be a NAND-TM program computing $F$ in $T(n)$ steps. We construct an oblivious NAND-TM program $P$ for computing $F$ as follows (see also Fig. 12.8).

1. On input $x$, $P$ will compute $T = T(|x|)$ and set up arrays $At\text{start}$ and $At\text{end}$ satisfying $At\text{start}[0] = 1$ and $At\text{start}[i] = 0$ for $i > 0$ and $At\text{end}[T - 1] = 1$ and $At\text{end}[i] = 0$ for all $i \neq T - 1$. We can do this because $T$ is a nice function. Note that since this computation does not depend on $x$ but only on its length, it is oblivious.

2. $P$ will also have a special array $\text{Marker}$ initialized to all zeroes.

3. The index variable of $P$ will change direction of movement to the right whenever $At\text{start}[i] = 1$ and to the left whenever $At\text{end}[i] = 1$.

4. The program $P$ simulates the execution of $P'$. However, if the MODANDJMP instruction in $P'$ attempts to move to the right when $P$ is moving left (or vice versa) then $P$ will set $\text{Marker}[i]$ to 1 and enter into a special “waiting mode”. In this mode $P$ will wait until the next time in which $\text{Marker}[i] = 1$ (at the next sweep) at which points $P$ zeroes $\text{Marker}[i]$ and continues with the simulation. In the worst case this will take $2T(n)$ steps (if $P$ has to go all the way from one end to the other and back again.)

5. We also modify $P$ to ensure it ends the computation after simulating exactly $T(n)$ steps of $P'$, adding “dummy steps” if $P'$ ends early.

We see that $P$ simulates the execution of $P'$ with an overhead of $O(T(n))$ steps of $P$ per one step of $P'$, hence completing the proof.
Theorem 12.11 implies Theorem 12.10. Indeed, if \( P \) is a \( k \)-line oblivious NAND-TM program computing \( F \) in time \( T(n) \) then for every \( n \) we can obtain a NAND-CIRC program of \( (k - 1) \cdot T(n) \) lines by simply making \( T(n) \) copies of \( P \) (dropping the final MODANDJMP line). In the \( j \)-th copy we replace all references of the form \( \text{foo}[i] \) to \( \text{foo}_{i,j} \) where \( i_j \) is the value of \( i \) in the \( j \)-th iteration.

12.6.2 Algorithmic transformation of NAND-TM to NAND

The proof of Theorem 12.10 is algorithmic, in the sense that the proof yields a polynomial-time algorithm that given a Turing Machine \( M \) and parameters \( T \) and \( n \), produces a circuit of \( O(T^2) \) gates that agrees with \( M \) on all inputs \( x \in \{0, 1\}^n \) (as long as \( M \) runs for less than \( T \) steps these inputs.) We record this fact in the following theorem, since it will be useful for us later on:

**Theorem 12.12 — Turing-machine to circuit compiler.** There is algorithm \( \text{UNROLL} \) such that for every Turing Machine \( M \) and numbers \( n, T \), \( \text{UNROLL}(M, T, n) \) runs for \( \text{poly}(T) \) steps and outputs a NAND circuit \( C \) with \( O(T^2) \) gates satisfying

\[
C(x) = M(x)
\]

for every \( x \in \{0, 1\}^n \) on which \( M \) halts within at most \( T \) steps.

**Proof.** We only sketch the proof since it follows by directly translating the proof of non-uniform-thm{.ref into an algorithm together with the simulation of Turing machines by NAND-TM programs (see also Fig. 12.9). Specifically, \( \text{UNROLL} \) does the following:

![Figure 12.9](image-url) We can transform a Turing Machine \( M \), input length parameter \( n \), and time bound \( T \) into an \( O(T^2) \) sized NAND circuit that agrees with \( M \) on all inputs \( x \in \{0, 1\}^n \) on which \( M \) halts in at most \( T \) steps. The transformation is obtained by first using the equivalence of Turing Machines and NAND-TM programs \( P \), then turning \( P \) into an equivalent oblivious NAND-TM program \( P' \) via Theorem 12.11, then “unrolling” \( O(T^2) \) iterations of the loop of \( P' \) to obtain an \( O(T^2) \) line NAND-CIRC program that agrees with \( P' \) on length \( n \) inputs, and finally translating this program into an equivalent circuit.
1. Transform the Turing Machine $M$ into an equivalent NAND-TM program $P$.

2. Transform the NAND-TM program $P$ into an equivalent oblivious program $P'$ following the proof of Theorem 12.11. The program $P'$ takes $T' = O(T^2)$ steps to simulate $T$ steps of $P$.

3. “Unroll the loop” of $P'$ by obtaining a NAND-CIRC program of $O(T')$ lines (or equivalently a NAND circuit with $O(T^2)$ gates) corresponding to the execution of $T'$ iterations of $P'$.

Reviewing the transformations described in Fig. 12.9, as well as solving the following two exercises is a great way to get more comfort with non-uniform complexity and in particular with $P/\text{poly}$ and its relation to $P$.

Solved Exercise 12.3 — Alternative characterization of $P$. Prove that for every $F : \{0, 1\}^* \rightarrow \{0, 1\}$, $F \in P$ if and only if there is a polynomial-time Turing Machine $M$ such that for every $n \in \mathbb{N}$, $M(1^n)$ outputs a description of an $n$ input circuit $C_n$ that computes the restriction $F \upharpoonright n$ of $F$ to inputs in $\{0, 1\}^n$.

Solution:

We start with the “if” direction. Suppose that there is a polynomial-time Turing Machine $M$ that on input $1^n$ outputs a circuit $C_n$ that computes $F \upharpoonright n$. Then the following is a polynomial-time Turing Machine $M'$ to compute $F$. On input $x \in \{0, 1\}^*$, $M'$ will:

1. Let $n = |x|$ and compute $C_n = M(1^n)$.
2. Return the evaluation of $C_n$ on $x$.

Since we can evaluate a Boolean circuit on an input in polynomial time, $M'$ runs in polynomial time and computes $F(x)$ on every input $x$.

For the “only if” direction, if $M'$ is a Turing Machine that computes $F$ in polynomial-time, then (applying the equivalence of Turing Machines and NAND-TM as well as Theorem 12.11) there is also an oblivious NAND-TM program $P$ that computes $F$ in time $p(n)$ for some polynomial $p$. We can now define $M$ to be the Turing Machine that on input $1^n$ outputs the NAND circuit obtained by “unrolling the loop” of $P$ for $p(n)$ iterations. The resulting NAND
circuit computes $F|_n$ and has $O(p(n))$ gates. It can also be transformed to a Boolean circuit with $O(p(n))$ AND/OR/NOT gates.

**Solved Exercise 12.4 — $P_{/\text{poly}}$ characterization by advice.** Let $F : \{0,1\}^* \rightarrow \{0,1\}$. Then $F \in P_{/\text{poly}}$ if and only if there exists a polynomial $p : \mathbb{N} \rightarrow \mathbb{N}$, a polynomial-time Turing Machine $M$ and a sequence $\{a_n\}_{n \in \mathbb{N}}$ of strings, such that for every $n \in \mathbb{N}$:

- $|a_n| \leq p(n)$
- For every $x \in \{0,1\}^n$, $M(a_n, x) = F(x)$.

**Solution:**

We only sketch the proof. For the “only if” direction, if $F \in P_{/\text{poly}}$ then we can use for $a_n$ simply the description of the corresponding circuit $C_n$ and for $M$ the program that computes in polynomial time the evaluation of a circuit on its input.

For the “if” direction, we can use the same “unrolling the loop” technique of Theorem 12.10 to show that if $P$ is a polynomial-time NAND-TM program, then for every $n \in \mathbb{N}$, the map $x \mapsto P(a_n, x)$ can be computed by a polynomial size NAND-CIRC program $Q_n$.

### 12.6.3 Can uniform algorithms simulate non uniform ones?

**Theorem 12.10** shows that every function in $\text{TIME}(T(n))$ is in $\text{SIZE}(\text{poly}(T(n)))$. One can ask if there is an inverse relation. Suppose that $F$ is such that $F|_n$ has a “short” NAND-CIRC program for every $n$. Can we say that it must be in $\text{TIME}(T(n))$ for some “small” $T$? The answer is an emphatic no. Not only is $P_{/\text{poly}}$ not contained in $P$, in fact $P_{/\text{poly}}$ contains functions that are uncomputable!

**Theorem 12.13 — $P_{/\text{poly}}$ contains uncomputable functions.** There exists an uncomputable function $F : \{0,1\}^* \rightarrow \{0,1\}$ such that $F \in P_{/\text{poly}}$.

**Proof Idea:**

Since $P_{/\text{poly}}$ corresponds to non uniform computation, a function $F$ is in $P_{/\text{poly}}$ if for every $n \in \mathbb{N}$, the restriction $F|_n$ to inputs of length $n$ has a small circuit/program, even if the circuits for different values of $n$ are completely different from one another. In particular, if $F$ has the property that for every equal-length inputs $x$ and $x'$, $F(x) = F(x')$ then this means that $F|_n$ is either the constant function zero or the constant function one for every $n \in \mathbb{N}$. Since the constant function has a (very!) small circuit, such a function $F$ will always be in $P_{/\text{poly}}$ (indeed even in smaller classes). Yet by a reduction from
the Halting problem, we can obtain a function with this property that is uncomputable.

\* 

**Proof of Theorem 12.13.** Consider the following “unary halting function” $UH : \{0, 1\}^* \rightarrow \{0, 1\}$ defined as follows. We let $S : \mathbb{N} \rightarrow \{0, 1\}^*$ be the function that on input $n \in \mathbb{N}$, outputs the string that corresponds to the binary representation of the number $n$ without the most significant 1 digit. Note that $S$ is onto. For every $x \in \{0, 1\}$, we define $UH(x) = HALTONZERO(S(|x|))$. That is, if $n$ is the length of $x$, then $UH(x) = 1$ if and only if the string $S(n)$ encodes a NAND-TM program that halts on the input 0.

$UH$ is uncomputable, since otherwise we could compute $HALTONZERO$ by transforming the input program $P$ into the integer $n$ such that $P = S(n)$ and then running $UH(1^n)$ (i.e., $UH$ on the string of $n$ ones). On the other hand, for every $n$, $UH_n(x)$ is either equal to 0 for all inputs $x$ or equal to 1 on all inputs $x$, and hence can be computed by a NAND-CIRC program of a constant number of lines.

\[\]

The issue here is of course uniformity. For a function $F : \{0, 1\}^* \rightarrow \{0, 1\}$, if $F$ is in $\text{TIme}(T(n))$ then we have a single algorithm that can compute $F|_n$ for every $n$. On the other hand, $F|_n$ might be in $\text{Size}(T(n))$ for every $n$ using a completely different algorithm for every input length. For this reason we typically use $P_{/\text{poly}}$ not as a model of efficient computation but rather as a way to model inefficient computation. For example, in cryptography people often define an encryption scheme to be secure if breaking it for a key of length $n$ requires more than a polynomial number of NAND lines. Since $P \subseteq P_{/\text{poly}}$, this in particular precludes a polynomial time algorithm for doing so, but there are technical reasons why working in a non-uniform model makes more sense in cryptography. It also allows to talk about security in non-asymptotic terms such as a scheme having “128 bits of security”.

While it can sometimes be a real issue, in many natural settings the difference between uniform and non-uniform computation does not seem to so important. In particular, in all the examples of problems not known to be in $P$ we discussed before: longest path, 3SAT, factoring, etc., these problems are also not known to be in $P_{/\text{poly}}$ either. Thus, for “natural” functions, if you pretend that $\text{TIme}(T(n))$ is roughly the same as $\text{Size}(T(n))$, you will be right more often than wrong.
12.6.4 Uniform vs. Nonuniform computation: A recap

To summarize, the two models of computation we have described so far are:

- **Uniform models**: Turing machines, NAND-TM programs, RAM machines, NAND-RAM programs, C/JavaScript/Python, etc. These models include loops and unbounded memory hence a single program can compute a function with unbounded input length.

- **Non-uniform models**: Boolean Circuits or straightline programs have no loops and can only compute finite functions. The time to execute them is exactly the number of lines or gates they contain.

For a function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ and some nice time bound $T : \mathbb{N} \rightarrow \mathbb{N}$, we know that:

- If $F$ is uniformly computable in time $T(n)$ then there is a sequence of circuits $C_1, C_2, \ldots$ where $C_n$ has $\text{poly}(T(n))$ gates and computes $F_{|n}$ (i.e., restriction of $F$ to $\{0, 1\}^n$) for every $n$.

- The reverse direction is not necessarily true - there are examples of functions $F : \{0, 1\}^n \rightarrow \{0, 1\}$ such that $F_{|n}$ can be computed by even a constant size circuit but $F$ is uncomputable.

This means that non uniform complexity is more useful to establish hardness of a function than its easiness.

---

**Lecture Recap**

- We can define the time complexity of a function using NAND-TM programs, and similarly to the notion of computability, this appears to capture the inherent complexity of the function.

- There are many natural problems that have polynomial-time algorithms, and other natural problems that we’d love to solve, but for which the best known algorithms are exponential.
• The definition of polynomial time, and hence the class $P$, is robust to the choice of model, whether it is Turing machines, NAND-TM, NAND-RAM, modern programming languages, and many other models.

• The time hierarchy theorem shows that there are some problems that can be solved in exponential, but not in polynomial time. However, we do not know if that is the case for the natural examples that we described in this lecture.

• By “unrolling the loop” we can show that every function computable in time $T(n)$ can be computed by a sequence of NAND-CIRC programs (one for every input length) each of size at most $\text{poly}(T(n))$.

12.7 EXERCISES

Exercise 12.1 — Equivalence of different definitions of $P$ and $\text{EXP}$. Prove that the classes $P$ and $\text{EXP}$ defined in Definition 12.2 are equal to $\cup_{c \in \{1, 2, 3, \ldots\}} \text{TIME}(n^c)$ and $\cup_{c \in \{1, 2, 3, \ldots\}} \text{TIME}(2^{nc})$ respectively. (If $S_1, S_2, S_3, \ldots$ is a collection of sets then the set $S = \cup_{c \in \{1, 2, 3, \ldots\}} S_c$ is the set of all elements $e$ such that there exists some $c \in \{1, 2, 3, \ldots\}$ where $e \in S_c$.)

Exercise 12.2 — Robustness to representation. Theorem 12.5 shows that the classes $P$ and $\text{EXP}$ are robust with respect to variations in the choice of the computational model. This exercise shows that these classes are also robust with respect to our choice of the representation of the input.

Specifically, let $F$ be a function mapping graphs to $\{0, 1\}$, and let $F', F'' : \{0, 1\}^* \rightarrow \{0, 1\}$ be the functions defined as follows. For every $x \in \{0, 1\}^*$:

• $F'(x) = 1$ iff $x$ represents a graph $G$ via the adjacency matrix representation such that $F(G) = 1$.

• $F''(x) = 1$ iff $x$ represents a graph $G$ via the adjacency list representation such that $F(G) = 1$.

Prove that $F' \in P$ iff $F'' \in P$.

More generally, for every function $F : \{0, 1\}^* \rightarrow \{0, 1\}$, the answer to the question of whether $F \in P$ (or whether $F \in \text{EXP}$) is unchanged by switching representations, as long as transforming one representation to the other can be done in polynomial time (which essentially holds for all reasonable representations).
Exercise 12.3 — Boolean functions. For every function \( F : \{0, 1\}^* \to \{0, 1\}^* \), define \( \text{Bool}(F) \) to be the function mapping \( \{0, 1\}^* \) to \( \{0, 1\} \) such that on input a (string representation of a) triple \((x, i, \sigma)\) with \( x \in \{0, 1\}^* \), \( i \in \mathbb{N} \) and \( \sigma \in \{0, 1\} \),

\[
\text{Bool}(F)(x, i, \sigma) = \begin{cases} 
F(x)_i & \sigma = 0, i < |F(x)| \\
1 & \sigma = 1, i < |F(x)| \\
0 & \text{otherwise}
\end{cases}
\] (12.5)

where \( F(x)_i \) is the \( i \)-th bit of the string \( F(x) \).

Prove that \( F \in \mathcal{P} \) if and only if \( \text{Bool}(F) \in \mathcal{P} \).

Exercise 12.4 — Composition of polynomial time. Prove that if \( F, G : \{0, 1\}^* \to \{0, 1\}^* \) are in \( \mathcal{P} \) then their composition \( F \circ G \), which is the function \( H \) s.t. \( H(x) = F(G(x)) \), is also in \( \mathcal{P} \).

Exercise 12.5 — Non composition of exponential time. Prove that there is some \( F, G : \{0, 1\}^* \to \{0, 1\}^* \) s.t. \( F, G \in \mathcal{E} \mathcal{X} \mathcal{P} \) but \( F \circ G \) is not in \( \mathcal{E} \mathcal{X} \mathcal{P} \).

Exercise 12.6 — Oblivious program. We say that a NAND-TM program \( P \) is oblivious if there is some functions \( T : \mathbb{N} \to \mathbb{N} \) and \( i : \mathbb{N} \times \mathbb{N} \to \mathbb{N} \) such that for every input \( x \) of length \( n \), it holds that:

* \( P \) halts when given input \( x \) after exactly \( T(n) \) steps.
* For \( t \in \{1, \ldots, T(n)\} \), after \( P \) executes the \( t \)-th step of the execution the value of the index \( i \) is equal to \( t(n, i) \). In particular this value does not depend on \( x \) but only on its length.\(^3\)

Let \( F : \{0, 1\}^* \to \{0, 1\}^* \) be such that there is some function \( m : \mathbb{N} \to \mathbb{N} \) satisfying \( |F(x)| = m(|x|) \) for every \( x \), and let \( P \) be a NAND-TM program that computes \( F \) in \( T(n) \) time for some nice \( T \). Then there is an oblivious NAND-TM program \( P' \) that computes \( F \) in time \( O(T^2(n) \log T(n)) \).

Exercise 12.7 Let \( \text{EDGE} : \{0, 1\}^* \to \{0, 1\} \) be the function such that on input a string representing a triple \((L, i, j)\), where \( L \) is the adjacency list representation of an \( n \) vertex graph \( G \), and \( i \) and \( j \) are numbers in \( [n] \), \( \text{EDGE}(L, i, j) = 1 \) if the edge \( \{i, j\} \) is present in the graph. \( \text{EDGE} \) outputs 0 on all other inputs.

1. Prove that \( \text{EDGE} \in \mathcal{P} \).
2. Let \( \text{PLANARMATRIX} : \{0, 1\}^* \to \{0, 1\} \) be the function that on input an adjacency matrix \( A \) outputs 1 if and only if the graph represented by \( A \) is planar (that is, can be drawn on the plane without edges crossing one another). For this question, you can use without proof the fact that \( \text{PLANARMATRIX} \in \mathcal{P} \). Prove that

\(^3\) An oblivious program \( P \) cannot compute functions whose output length is not a function of the input length, though this is not a real restriction, as we can always embed variable output functions in fixed length ones using some special “end of output” marker.
PLANARLIST ∈ P where PLANARLIST : \{0, 1\}^* → \{0, 1\} is the function that on input an adjacency list L outputs 1 if and only if L represents a planar graph.

**Exercise 12.8 — Evaluate NAND circuits.** Let NANDVAL : \{0, 1\}^* → \{0, 1\} be the function such that for every string representing a pair (Q, x) where Q is an n-input 1-output NAND-CIRC (not NAND-TM!) program and x ∈ \{0, 1\}^n, NANDVAL(Q, x) = Q(x). On all other inputs NANDVAL outputs 0.

Prove that NANDVAL ∈ P.

**Exercise 12.9 — Find hard function.** Let NANDHARD : \{0, 1\}^* → \{0, 1\} be the function such that on input a string representing a pair (f, s) where

\begin{itemize}
  \item f ∈ \{0, 1\}^{2^n} for some n ∈ \N
  \item s ∈ \N
\end{itemize}

NANDHARD(f, s) = 1 if there is no NAND-CIRC program Q of at most s lines that computes the function F : \{0, 1\}^n → \{0, 1\} whose truth table is the string f. That is, NANDHARD(f, s) = 1 if for every NAND-CIRC program Q of at most s lines, there exists some x ∈ \{0, 1\}^n such that Q(x) ≠ f_x where f_x denote the x-the coordinate of f, using the binary representation to identify \{0, 1\}^n with the numbers \{0, \ldots, 2^n - 1\}.

1. Prove that NANDHARD ∈ EXP.

2. (Challenge) Prove that there is an algorithm FINDHARD such that if n is sufficiently large, then FINDHARD(1^n) runs in time 2^{2^{100n}} and outputs a string f ∈ \{0, 1\}^{2^n} that is the truth table of a function that is not contained in SIZE(2^n/(1000n)). (In other words, if f is the string output by FINDHARD(1^n) then if we let F : \{0, 1\}^n → \{0, 1\} be the function such that F(x) outputs the x-th coordinate of f, then F \notin SIZE(2^n/(1000n)).

**Exercise 12.10** Suppose that you are in charge of scheduling courses in computer science in University X. In University X, computer science students wake up late, and have to work on their startups in the afternoon, and take long weekends with their investors. So you only have two possible slots: you can schedule a course either Monday-Wednesday 11am-1pm or Tuesday-Thursday 11am-1pm.

Let SCHEDULE : \{0, 1\}^* → \{0, 1\} be the function that takes as input a list of courses L and a list of conflicts C (i.e., list of pairs of courses

\[^4\text{Hint: Use Item 1, the existence of functions requiring exponentially hard NAND programs, and the fact that there are only finitely many functions mapping \{0, 1\}^n to \{0, 1\}.}
that cannot share the same time slot) and outputs 1 if and only if there is a “conflict free” scheduling of the courses in $L$, where no pair in $C$ is scheduled in the same time slot.

More precisely, the list $L$ is a list of strings $(c_0, \ldots, c_{n-1})$ and the list $C$ is a list of pairs of the form $(c_i, c_j)$. $SCHEDULE(L, C) = 1$ if and only if there exists partition of $c_0, \ldots, c_{n-1}$ into two parts so that there is no pair $(c_i, c_j) \in C$ such that both $c_i$ and $c_j$ are in the same part.

Prove that $SCHEDULE \in \mathbf{P}$. As usual, you do not have to provide the full code to show that this is the case, and can describe operations as a high level, as well as appeal to any data structures or other results mentioned in the book or in lecture. Note that to show that a function $F$ is in $\mathbf{P}$ you need to both (1) present an algorithm $A$ that computes $F$ in polynomial time, (2) prove that $A$ does indeed run in polynomial time, and does indeed compute the correct answer.

Try to think whether or not your algorithm extends to the case where there are three possible time slots.

12.8 BIBLIOGRAPHICAL NOTES

Because we are interested in the maximum number of steps for inputs of a given length, running-time as we defined it is often known as worst case complexity. The minimum number of steps (or “best case” complexity) to compute a function on length $n$ inputs is typically not a meaningful quantity since essentially every natural problem will have some trivially easy instances. However, the average case complexity (i.e., complexity on a “typical” or “random” input) is an interesting concept which we’ll return to when we discuss cryptography. That said, worst-case complexity is the most standard and basic of the complexity measures, and will be our focus in most of this book.

Some lower bounds for single-tape Turing machines are given in [Maa85].

For defining efficiency in the $\lambda$ calculus, one needs to be careful about the order of application of the reduction steps, which can matter for computational efficiency, see for example this paper.

The notation $\mathbf{P}_{/poly}$ is used for historical reasons. It was introduced by Karp and Lipton, who considered this class as corresponding to functions that can be computed by polynomial-time Turing Machines that are given for any input length $n$ an advice string of length polynomial in $n$. 
Consider some of the problems we have encountered in Chapter 11:

- The 3SAT problem: deciding whether a given 3CNF formula has a satisfying assignment.
- Finding the longest path in a graph.
- Finding the maximum cut in a graph.
- Solving quadratic equations over $n$ variables $x_0, x_1, ..., x_{n-1} \in \mathbb{R}$.

All of these problems have the following properties:

- These are important problems, and people have spent significant effort on trying to find better algorithms for them.
- Each one of these is a search problem, whereby we search for a solution that is “good” in some easy to define sense (e.g., a long path, a satisfying assignment, etc.).
- Each of these problems has a trivial exponential time algorithm that involve enumerating all possible solutions.
- At the moment, for all these problems the best known algorithm is not much faster than the trivial one in the worst case.

In this chapter and in Chapter 14 we will see that, despite their apparent differences, we can relate the computational complexity of these and many other problems. In fact, it turns out that the problem above are computationally equivalent, in the sense that solving one of them immediately implies solving the others. This phenomenon, known as \textit{NP completeness}, is one of the surprising discoveries of theoretical computer science, and we will see that it has far-reaching ramifications.

In this chapter we will see that for each one of the problems of finding a longest path in a graph, solving quadratic equations, and finding
the maximum cut, if there exists a polynomial-time algorithm for this problem then there exists a polynomial-time algorithm for the 3SAT problem as well. In other words, we will reduce the task of solving 3SAT to each one of the above tasks. Another way to interpret these results is that if there does not exist a polynomial-time algorithm for 3SAT then there does not exist a polynomial-time algorithm for these other problems as well. In Chapter 14 we will see evidence (though no proof!) that all of the above problem do not have polynomial-time algorithms and hence are inherently intractable.

**Decision problems.** For reasons of technical conditions rather than anything substantial, we will concern ourselves with decision problems (i.e., Yes/No questions) or in other words Boolean (i.e., one-bit output) functions. We model the problems above as functions mapping \(\{0, 1\}^*\) to \(\{0, 1\}\) in the following way:

The 3SAT problem can be phrased as the function \(3SAT : \{0, 1\}^* \rightarrow \{0, 1\}\) that takes as input a 3CNF formula \(\varphi\) (i.e., a formula of the form \(C_0 \land \cdots \land C_{m-1}\) where each \(C_i\) is the OR of three variables or their negation) and maps \(\varphi\) to 1 if there exists some assignment to the variables of \(\varphi\) that causes it to evaluate to true, and to 0 otherwise. For example

\[
3SAT(\left(\neg x_0 \land x_1 \land x_2\right) \lor \left(x_1 \lor x_2 \lor \neg x_3\right) \lor \left(x_0 \land \neg x_2 \land x_3\right)) = 1 \quad (13.1)
\]

since the assignment \(x = 1101\) satisfies the input formula. In the above we assume some representation of formulas as strings, and define the function to output 0 if its input is not a valid representation; we use the same convention for all the other functions below.
The quadratic equations problem corresponds to the function $\text{QUADEQ} : \{0, 1\}^* \rightarrow \{0, 1\}$ that maps a set of quadratic equations $E$ to 1 if there is an assignment $x$ that satisfies all equations, and to 0 otherwise.

The longest path problem corresponds to the function $\text{LONGPATH} : \{0, 1\}^* \rightarrow \{0, 1\}$ that maps a graph $G$ and a number $k$ to 1 if there is a simple path in $G$ of length at least $k$, and maps $(G, k)$ to 0 otherwise. The longest path problem is a generalization of the well-known Hamiltonian Path Problem of determining whether a path of length $n$ exists in a given $n$ vertex graph.

The maximum cut problem corresponds to the function $\text{MAXCUT} : \{0, 1\}^* \rightarrow \{0, 1\}$ that maps a graph $G$ and a number $k$ to 1 if there is a cut in $G$ that cuts at least $k$ edges, and maps $(G, k)$ to 0 otherwise.

All of the problems above are in EXP but it is not known whether or not they are in P. However, we will see in this chapter that if either $\text{QUADEQ}$, $\text{LONGPATH}$ or $\text{MAXCUT}$ are in P, then so is 3SAT.

13.1 REDUCTIONS

Suppose that $F, G : \{0, 1\}^* \rightarrow \{0, 1\}$ are two functions. A reduction from $F$ to $G$ is a way to show that $F$ is “no harder” than $G$, in the sense that a polynomial-time algorithm for $G$ implies a polynomial-time algorithm for $F$. The formal definition is as follows:

**Definition 13.1 — Reductions.** Let $F, G : \{0, 1\}^* \rightarrow \{0, 1\}^*$. We say that $F$ reduces to $G$, denoted by $F \leq_p G$ if there is a polynomial-time computable $R : \{0, 1\}^* \rightarrow \{0, 1\}^*$ such that for every $x \in \{0, 1\}^*$, $F(x) = G(R(x))$.

We say that $F$ and $G$ have equivalent complexity if $F \leq_p G$ and $G \leq_p F$.

The following exercise justifies our intuition that $F \leq_p G$ signifies that “$F$ is no harder than $G$.

**Solved Exercise 13.1 — Reductions and P.** Prove that if $F \leq_p G$ and $G \in \text{P}$ then $F \in \text{P}$.
Solution:

Suppose there was an algorithm $B$ that compute $F$ in time $p(n)$ where $p$ is its input size. Then, (13.2) directly gives an algorithm $A$ to compute $F$ (see Fig. 13.2). Indeed, on input $x \in \{0, 1\}^*$, Algorithm $A$ will run the polynomial-time reduction $R$ to obtain $y = R(x)$ and then return $B(y)$. By (13.2), $G(R(x)) = F(x)$ and hence Algorithm $A$ will indeed compute $F$.

We now show that $A$ runs in polynomial time. By assumption, $R$ can be computed in time $q(n)$ for some polynomial $q$. In particular, this means that $|y| \leq q(|x|)$ (as just writing down $y$ takes $|y|$ steps). This, computing $B(y)$ will take at most $p(|y|) \leq p(q(|x|))$ steps. Thus the total running time of $A$ on inputs of length $n$ is at most the time to compute $y$, which is bounded by $q(n)$, and the time to compute $B(y)$, which is bounded by $p(q(n))$, and since the composition of two polynomials is a polynomial, $A$ runs in polynomial time.

A reduction from $F$ to $G$ can be used for two purposes:

- If we already know an algorithm for $G$ and $F \leq_p G$ then we can use the reduction to obtain an algorithm for $F$. This is a widely used tool in algorithm design. The “quicksort” algorithm reduces the task of sorting to the task of partitioning an array to the elements smaller and bigger than some “pivot” element. Dijkstra’s shortest path algorithm reduces the task of finding the shortest path between two elements to the task of implementing a priority queue data structure.

- If we have proven (or have evidence) that there exists no polynomial-time algorithm for $F$ and $F \leq_p G$ then the existence of this reduction allows us to concludes that there exists no polynomial-time algorithm for $G$. This is the “if pigs could whistle then horses could fly” interpretation we’ve seen in Section 8.4. We show that if there was an hypothetical efficient algorithm for $G$ (a “whistling pig”) then since $F \leq_p G$ then there would be an efficient algorithm for $F$ (a “flying horse”).

In this book we will often use reductions for the second purpose, although the lines between the two is sometimes blurry (see the bibliographical notes).

Since we think of $F \leq_p G$ as saying that (as far as polynomial-time computation is concerned) $F$ is “easier or equal in difficulty to” $G$, we would expect that if $F \leq_p G$ and $G \leq_p H$, then it would hold that $F \leq_p H$. Indeed this is the case:
**Lemma 13.2** For every $F, G, H : \{0, 1\}^* \to \{0, 1\}$, if $F \leq_p G$ and $G \leq_p H$ then $F \leq_p H$.

We leave the proof of Lemma 13.2 as Exercise 13.1. Pausing now and doing this exercise is an excellent way to verify that you understood the definition of reductions.

The most crucial difference between the notion in Definition 13.1 and the reductions we saw in the context of uncomputability (e.g., in Section 8.4) is that for relating time complexity of problems, we need the reduction to be computable in *polynomial time*, as opposed to merely computable. Definition 13.1 also restricts reductions to have a very specific format. That is, to show that $F \leq_p G$, rather than allowing a general algorithm for $F$ that uses a “magic box” that computes $G$, we only allow an algorithm that computes $F(x)$ by outputting $G(R(x))$. This restricted form is convenient for us, but people have defined and used more general reductions as well (see Section 13.8).

In this chapter we use reductions to relate the computational complexity of the problems mentioned above: 3SAT, Quadratic Equations, Maximum Cut, and Longest Path, as well as a few others. We will reduce 3SAT to the latter problems, demonstrating that solving any one of them efficiently will result in an efficient algorithm for 3SAT. In Chapter 14 we show the other direction: reducing each one of these problems to 3SAT in one fell swoop.

### 13.2 REDUCING 3SAT TO ZERO ONE EQUATIONS

We will now show our first example of a reduction. The *Zero-One Linear Equations problem* corresponds to the function $01EQ : \{0, 1\}^* \to \{0, 1\}$ whose input is a collection $E$ of linear equations in variables $x_0, \ldots, x_{n-1}$, and the output is 1 iff there is an assignment $x \in \{0, 1\}^n$ of 0/1 values to the variables that satisfies all the equations. For example, if the input $E$ is a string encoding the set of equations

$$
\begin{align*}
x_0 + x_1 + x_2 &= 2 \\
x_0 + x_2 &= 1 \\
x_1 + x_2 &= 2
\end{align*}
$$

then $01EQ(E) = 1$ since the assignment $x = 011$ satisfies all three equations. We specifically restrict attention to linear equations in variables $x_0, \ldots, x_{n-1}$ in which every equation has the form $\sum_{i \in S} x_i = b$ where $S \subseteq [n]$ and $b \in \mathbb{N}$.

If we asked the question of whether there is a solution $x \in \mathbb{R}^n$ of real numbers to $E$, then this can be solved using the famous Gaussian

---

1. If you are familiar with matrix notation you may note that such equations can be written as $Ax = b$ where $A$ is an $m \times n$ matrix with entries in 0/1 and $b \in \mathbb{N}^m$. 

elimination algorithm in polynomial time. However, there is no known efficient algorithm to solve 01EQ. Indeed, such an algorithm would imply an algorithm for 3SAT as shown by the following theorem:

**Theorem 13.3 — Hardness of 01EQ.** $3SAT \leq_p 01EQ$

Proof Idea:
A constraint $x_2 \lor x_5 \lor x_7$ can be written as $x_2 + (1 - x_5) + x_7 \geq 1$. This is a linear inequality but since the sum on the left-hand side is at most three, we can also turn it into an equality by adding two new variables $y, z$ and writing it as $x_2 + (1 - x_5) + x_7 + y + z = 3$. (We will use fresh such variables $y, z$ for every constraint.) Finally, for every variable $x_i$ we can add a variable $x_i'$ corresponding to its negation by adding the equation $x_i + x_i' = 1$, hence mapping the original constraint $x_2 \lor x_5 \lor x_7$ to $x_2 + x_5' + x_7 + y + z = 3$.

*  

Proof of Theorem 13.3. To prove the theorem we need to:

1. Describe an algorithm $R$ for mapping an input $\varphi$ for 3SAT into an input $E$ for 01EQ.

2. Prove that the algorithm runs in polynomial time and that $01EQ(R(\varphi)) = 3SAT(\varphi)$ for every 3CNF formula $\varphi$.

We proceed to do just that.

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### 13.3 Quadratic Equations

Recall that in the quadratic equation problem, the input is a list of $n$-variate polynomials $p_0, \ldots, p_{m-1} : \mathbb{R}^n \rightarrow \mathbb{R}$ that are all of degree at most two (i.e., they are quadratic) and with integer coefficients. (The latter condition is for convenience and can be achieved by scaling.) The task is to find out whether there is a solution $x \in \mathbb{R}^n$ to the equations $p_0(x) = 0, p_1(x) = 0, \ldots, p_{m-1}(x) = 0$. 

For example, the following is a set of quadratic equations over the variables $x_0, x_1, x_2$:

\[
\begin{align*}
    x_0^2 - x_0 &= 0 \\
    x_1^2 - x_1 &= 0 \\
    x_2^2 - x_2 &= 0 \\
    1 - x_0 - x_1 + x_0 x_1 &= 0
\end{align*}
\] (13.4)

You can verify that $x \in \mathbb{R}^3$ satisfies this set of equations if and only if $x \in \{0, 1\}^3$ and $x_0 \lor x_1 = 1$.

We will show how to reduce 3SAT to the problem of Quadratic Equations.

**Theorem 13.4 — Hardness of quadratic equations.**

\[3SAT \leq_p QUADEQ\] (13.5)

where 3SAT is the function that maps a 3SAT formula $\varphi$ to 1 if it is satisfiable and to 0 otherwise, and QUADEQ is the function that maps a set $E$ of quadratic equations over $\{0, 1\}^n$ to 1 if it has a solution and to 0 otherwise.

**Proof Idea:**

At the end of the day, a 3SAT formula can be thought of as a list of equations on some variables $x_0, \ldots, x_{n-1}$. Namely, the equations are that each of the $x_i$'s should be equal to either 0 or 1, and that the variables should satisfy some set of constraints which corresponds to the OR of three variables or their negation. To show that $3SAT \leq_p QUADEQ$ we need to give a polynomial-time reduction that maps a 3SAT formula $\varphi$ into a set of quadratic equations $E$ such that $E$ has a solution if and only if $\varphi$ is satisfiable. The idea is that we can transform a 3SAT formula $\varphi$ first to a set of cubic equations by mapping every constraint of the form $(x_1 \lor x_{15} \lor x_{24})$ into an equation of the form $(1-x_1) x_{15} (1-x_{24}) = 0$. We can then turn this into a quadratic equation by mapping any cubic equation of the form $x_i x_j x_k = 0$ into the two quadratic equations $y_{i,j} = x_i x_j$ and $y_{i,j} x_k = 0$.

*Proof of Theorem 13.4.* To prove Theorem 13.4 we need to give a polynomial-time transformation of every 3SAT formula $\varphi$ into a set of quadratic equations $E$, and prove that $3SAT(\varphi) = QUADEQ(E)$.

We now describe the transformation of a formula $\varphi$ to equations $E$ and show the completeness and soundness conditions. Recall that a 3SAT formula $\varphi$ is a formula such as $(x_{17} \lor \overline{x}_{101} \lor x_{57}) \land (x_{18} \lor \overline{x}_{19} \lor \overline{x}_{101}) \land \ldots$. That is, $\varphi$ is composed of the AND of $m$ 3SAT clauses
where a 3SAT clause is the OR of three variables or their negation. A quadratic equations instance $E$ is composed of a list of equations, each of involving a sum of variables or their products, such as $x_{19}x_{52} - x_{12} + 2x_{33} = 2$, etc. We will include the constraints $x_i^2 - x_i = 0$ for every $i \in [n]$ in our equations, which means that we can restrict attention to assignments where $x_i \in \{0, 1\}$ for every $i$.

There is a natural way to map a 3SAT instance into a set of cubic equations $E'$, and that is to map a clause such as $(x_{17} \lor x_{101} \lor x_{57})$ (which is equivalent to the negation of $x_{17} \land x_{101} \land x_{57}$) to the equation $(1 - x_{17})x_{101}(1 - x_{57}) = 0$. Therefore, we can map a formula $\varphi$ with $n$ variables $m$ clauses into a set $E'$ of $m + n$ cubic equations on $n$ variables (that is, one equation per each clause, plus one equation of the form $x_i^2 - x_i = 0$ for each variable to ensure that its value is in $\{0, 1\}$) such that every assignment $a \in \{0, 1\}^n$ to the $n$ variables satisfies the original formula if and only if it satisfies the equations of $E'$.

To make the equations quadratic we introduce for every two distinct $i, j \in [n]$ a variable $y_{i,j}$ and include the constraint $y_{i,j} - x_i x_j = 0$ in the equations. This is a quadratic equation that ensures that $y_{i,j} = x_i x_j$ for every such $i, j \in [n]$. Now we can turn any cubic equation in the $x$’s into a quadratic equation in the $x$ and $y$ variables. For example, we can “open up the parentheses” of an equation such as $(1 - x_{17})x_{101}(1 - x_{57}) = 0$ to $x_{101} - x_{17}x_{101} - x_{101}x_{57} + x_{17}x_{101}x_{57} = 0$. We can now replace the cubic term $x_{17}x_{101}x_{57}$ with the quadratic term $y_{17,101}x_{57}$. This can be done for every cubic equation in the same way, replacing any cubic term $x_i x_j x_k$ with the term $y_{i,j} x_k$. The end result will be a set of $m + n + \binom{n}{2}$ equations (one equation per clause, one equation per variable to ensure $x_i^2 - x_i = 0$, and one equation per pair $i, j$ to ensure $y_{i,j} = x_i x_j = 0$) on the $n + \binom{n}{2}$ variables $x_0, \ldots, x_{n-1}$ and $y_{i,j}$ for all pairs of distinct variables $i, j$.

To complete the proof we need to show that if we transform $\varphi$ to $E$ in this way then the 3SAT formula $\varphi$ is satisfiable if and only if the equations $E'$ have a solution. This is essentially immediate from the construction, but as this is our first reduction, we spell this out fully:

- **Completeness:** We claim that if $\varphi$ is satisfiable then the equations $E$ have a solution. To prove this we need to show how to transform a satisfying assignment $a \in \{0, 1\}^n$ to the variables of $\varphi$ (that is, $a_i$ is the value assigned to $x_i$) to a solution to the variables of $E$.

  Specifically, if $a \in \{0, 1\}^n$ is such an assignment then by design $a$ satisfies all the cubic equations $E'$ that we constructed above. But then, if we assign to the $n + \binom{n}{2}$ variables the values $a_0, \ldots, a_{n-1}$ and $\{a_i, a_j\}$ for all $\{i, j\} \subseteq [n]$ then by construction this will satisfy all the quadratic equations of $E$ as well.
• **Soundness:** We claim that if the equations $E$ have a solution then $\varphi$ is satisfiable. Indeed, suppose that $z \in \mathbb{R}^{n+\binom{n}{2}}$ is a solution to the equations $E$. A priori $z$ could be any vector of $n + \binom{n}{2}$ numbers, but the fact that $E$ contains the equations $x_i^2 - x_i = 0$ and $y_{i,j} - x_i x_j = 0$ means that if $z$ satisfies these equations then the values it assigns to $x_i$ must be in $\{0, 1\}$ for every $i$, and the value it assigns to $y_{i,j}$ must be $x_i x_j$ for every $\{i, j\} \subseteq [n]$. Therefore by the way we constructed our equations, the value assigned $x$ must be a solution of the original cubic equations $E'$ and hence also of the original formula $\varphi$, which in particular implies $\varphi$ is satisfiable.

This reduction can be easily implemented in about a dozen lines of Python or any other programming language, see Fig. 13.3.

13.4 THE INDEPENDENT SET PROBLEM

For a graph $G = (V, E)$, an independent set (also known as a stable set) is a subset $S \subseteq V$ such that there are no edges with both endpoints in $S$ (in other words, $E(S, S) = \emptyset$). Every “singleton” (set consisting of a single vertex) is trivially an independent set, but finding larger independent sets can be challenging. The **maximum independent set** problem (henceforth simply “independent set”) is the task of finding the largest independent set in the graph. The independent set problem is naturally related to scheduling problems: if we put an edge between two conflicting tasks, then an independent set corresponds to a set of tasks that can all be scheduled together without conflicts. But it also arises in very different settings, including trying to find structure in protein-protein interaction graphs.

To phrase independent set as a decision problem, we think of it as a function $ISET : \{0, 1\}^* \to \{0, 1\}$ that on input a graph $G$ and a number $k$ outputs $1$ if and only if the graph $G$ contains an independent set of size at least $k$. We will now reduce 3SAT to Independent set.

**Theorem 13.5 — Hardness of Independent Set.** $3SAT \leq_p ISET$.

**Proof Idea:**

The idea is that finding a satisfying assignment to a 3SAT formula corresponds to satisfying many local constraints without creating any conflicts. One can think of “$x_{17} = 0$” and “$x_{17} = 1$” as two conflicting events, and of the constraints $x_{17} \lor x_5 \lor x_9$ as creating a conflict between the events “$x_{17} = 0$”, “$x_5 = 1$” and “$x_9 = 0$”, saying that these three cannot simultaneously co-occur. Using these ideas, we can we can think of solving a 3SAT problem as trying to schedule non conflicting events, though the devil is, as usual, in the details.
Proof of Theorem 13.5. Given a 3SAT formula \( \varphi \) on \( n \) variables and with \( m \) clauses, we will create a graph \( G \) with \( 3m \) vertices as follows: (see Fig. 13.4 for an example)

- A clause \( C \) in \( \varphi \) has the form \( C = y \lor y' \lor y'' \) where \( y, y', y'' \) are literals (variables or their negation). For each such clause \( C \), we will add three vertices to \( G \), and label them \((C, y)\), \((C, y')\), and \((C, y'')\) respectively. We will also add the three edges between all pairs of these vertices, so they form a triangle. Since there are \( m \) clauses in \( \varphi \), the graph \( G \) will have \( 3m \) vertices.

- In addition to the above edges, we also add an edge between every pair vertices of the form \((C, y)\) and \((C', y')\) where \( y \) and \( y' \) are conflicting literals. That is, we add an edge between \((C, y)\) and \((C, y')\) if there is an \( i \) such that \( y = x_i \) and \( y' = \overline{x_i} \) or vice versa.

The above construction of \( G \) based on \( \varphi \) can clearly be carried out in polynomial time. Hence to prove the theorem we need to show that \( \varphi \) is satisfiable if and only if \( G \) contains an independent set of \( m \) vertices. We now show both directions of this equivalence.

Part 1: Completeness. The “completeness” direction is to show that if \( \varphi \) has a satisfying assignment \( x^* \), then \( G \) has an independent set \( S^* \) of \( m \) vertices. Let us now show this.

Indeed, suppose that \( \varphi \) has a satisfying assignment \( x^* \in \{0, 1\}^n \). Then for every clause \( C = y \lor y' \lor y'' \) of \( \varphi \), one of the literals \( y, y', y'' \) must evaluate to true under the assignment \( x^* \) (as otherwise it would not satisfy \( \varphi \)). We let \( S \) be a set of \( m \) vertices that is obtained by choosing for every clause \( C \) one vertex of the form \((C, y)\) such that \( y \) evaluates to true under \( x^* \). (If there is more than one such vertex for the same \( C \), we arbitrarily choose one of them.)

We claim that \( S \) is an independent set. Indeed, suppose otherwise that there was a pair of vertices \((C, y)\) and \((C', y')\) in \( S \) that have an edge between them. Since we picked one vertex out of each triangle corresponding to a clause, it must be that \( C \neq C' \). Hence the only way that there is an edge between \((C, y)\) and \((C, y')\) is if \( y \) and \( y' \) are conflicting literals (i.e. \( y = x_i \) and \( y' = \overline{x_i} \) for some \( i \)). But that would that they can’t both evaluate to true under the assignment \( x^* \), which contradicts the way we constructed the set \( S \). This completes the proof of the completeness condition.

Part 2: Soundness. The “soundness” direction is to show that if \( G \) has an independent set \( S^* \) of \( m \) vertices, then \( \varphi \) has a satisfying assignment \( x^* \in \{0, 1\}^n \). Let us now show this.
Indeed, suppose that \( G \) has an independent set \( S^* \) with \( m \) vertices. We will define an assignment \( x^* \in \{0,1\}^n \) for the variables of \( \varphi \) as follows. For every \( i \in [n] \), we set \( x^*_i \) according to the following rules:

- If \( S^* \) contains a vertex of the form \((C, x_i)\) then we set \( x^*_i = 1 \).
- If \( S^* \) contains a vertex of the form \((C, \neg x_i)\) then we set \( x^*_i = 0 \).
- If \( S^* \) does not contain a vertex of either of these forms, then it does not matter which value we give to \( x^*_i \), but for concreteness we’ll set \( x^*_i = 0 \).

The first observation is that \( x^* \) is indeed well defined, in the sense that the rules above do not conflict with one another, and ask to set \( x^*_i \) to be both 0 and 1. This follows from the fact that \( S^* \) is an independent set and hence if it contains a vertex of the form \((C, x_i)\) then it cannot contain a vertex of the form \((C, \neg x_i)\).

We now claim that \( x^* \) is a satisfying assignment for \( \varphi \). Indeed, since \( S^* \) is an independent set, it cannot have more than one vertex inside each one of the \( m \) triangles \((C, y), (C, y'), (C, y'')\) corresponding to a clause of \( \varphi \). Hence since \( |S^*| = m \), it must have exactly one vertex in each such triangle. For every clause \( C \) of \( \varphi \), if \((C, y)\) is the vertex in \( S^* \) in the triangle corresponding to \( C \), then by the way we defined \( x^* \), the literal \( y \) must evaluate to true, which means that \( x^* \) satisfies this clause. Therefore \( x^* \) satisfies all clauses of \( \varphi \), which is the definition of a satisfying assignment.

This completes the proof of Theorem 13.5

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**Solved Exercise 13.2 — Clique is equivalent to independent set.** The maximum clique problem corresponds to the function \( \text{CLIQUE} : \{0,1\}^* \to \{0,1\} \) such that for a graph \( G \) and a number \( k \), \( \text{CLIQUE}(G, k) = 1 \) iff there is a \( S \) subset of \( k \) vertices such that for every distinct \( u, v \in S \), the edge \( u, v \) is in \( G \). Such a set is known as a clique.

Prove that \( \text{CLIQUE} \leq_p \text{ISET} \) and \( \text{ISET} \leq_p \text{CLIQUE} \).

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**Solution:**

If \( G = (V, E) \) is a graph, we denote by \( \overline{G} \) its complement which is the graph on the same vertices \( V \) and such that for every distinct \( u, v \in V \), the edge \( \{u, v\} \) is present in \( \overline{G} \) if and only if this edge is not present in \( G \).

This means that for every set \( S \), \( S \) is an independent set in \( G \) if and only if \( S \) is a clique in \( \overline{S} \). Therefore for every \( k \), \( \text{ISET}(G, k) = \text{CLIQUE}(\overline{G}, k) \). Since the map \( G \mapsto \overline{G} \) can be computed efficiently,
this yields a reduction \( ISET \leq_p CLIQUE \). Moreover, since \( \overline{G} = G \) this yields a reduction in the other direction as well.

13.5 REDUCING INDEPENDENT SET TO MAXIMUM CUT

**Theorem 13.6 — Hardness of Max Cut.** \( ISET \leq_p MAXCUT \)

**Proof Idea:**
We will map a graph \( G \) into a graph \( H \) such that a large independent set in \( G \) becomes a partition cutting many edges in \( H \). We can think of a cut in \( H \) as coloring each vertex either “blue” or “red”. We will add a special “source” vertex \( s^* \), connect it to all other vertices, and assume without loss of generality that it is colored blue. Hence the more vertices we color red, the more edges from \( s^* \) we cut. Now, for every edge \( u, v \) in the original graph \( G \) we will add a special “gadget” which will be a small subgraph that involves \( u, v, \) the source \( s^* \), and two other additional vertices. We design the gadget in a way so that if the red vertices are not an independent set in \( G \) then the corresponding cut in \( H \) will be “penalized” in the sense that it would not cut as many edges. Once we set for ourselves this objective, it is not hard to find a gadget that achieves it— see the proof below.

**Proof of Theorem 13.6.** We will transform a graph \( G \) of \( n \) vertices and \( m \) edges into a graph \( H \) of \( n + 1 + 2m \) vertices and \( n + 5m \) edges in the following way: the graph \( H \) will contain all vertices of \( G \) (though not the edges between them!) and in addition to that will contain:

* A special vertex \( s^* \) that is connected to all the vertices of \( G \)
* For every edge \( e = \{u, v\} \in E(G) \), two vertices \( e_0, e_1 \) such that \( e_0 \) is connected to \( u \) and \( e_1 \) is connected to \( v \), and moreover we add the edges \( \{e_0, e_1\}, \{e_0, s^*\}, \{e_1, s^*\} \) to \( H \).

Theorem 13.6 will follow by showing that \( G \) contains an independent set of size at least \( k \) if and only if \( H \) has a cut cutting at least \( k + 4m \) edges. We now prove both directions of this equivalence:

**Part 1: Completeness.** If \( I \) is an independent \( k \)-sized set in \( G \), then we can define \( S \) to be a cut in \( H \) of the following form: we let \( S \) contain all the vertices of \( I \) and for every edge \( e = \{u, v\} \in E(G) \), if \( u \in I \) and \( v \notin I \) then we add \( e_0 \) to \( S \); if \( u \notin I \) and \( v \in I \) then we add \( e_1 \) to \( S \); and if \( u \notin I \) and \( v \notin I \) then we add both \( e_0 \) and \( e_1 \) to \( S \). (We don’t need to worry about the case that both \( u \) and \( v \) are in \( I \) since it is an independent set.) We can verify that in all cases the number of edges from \( S \) to its complement in the gadget corresponding to \( e \) will be four
(see Fig. 13.5). Since $s^*$ is not in $S$, we also have $k$ edges from $s^*$ to $I$, for a total of $k + 4m$ edges.

**Part 2: Soundness.** Suppose that $S$ is a cut in $H$ that cuts at least $C = k + 4m$ edges. We can assume that $s^*$ is not in $S$ (otherwise we can “flip” $S$ to its complement $\overline{S}$, since this does not change the size of the cut). Now let $I$ be the set of vertices in $S$ that correspond to the original vertices of $G$. If $I$ was an independent set of size $k$ then would be done. This might not always be the case but we will see that if $I$ is not an independent set then its also larger than $k$. Specifically, we define $m_{in} = |E(I, I)|$ be the set of edges in $G$ that are contained in $I$ and let $m_{out} = m - m_{in}$ (i.e., if $I$ is an independent set then $m_{in} = 0$ and $m_{out} = m$). By the properties of our gadget we know that for every edge $\{u, v\}$ of $G$, we can cut at most three edges when both $u$ and $v$ are in $S$, and at most four edges otherwise. Hence the number $C$ of edges cut by $S$ satisfies $C \leq |I| + 3m_{in} + 4m_{out} = |I| + 3m_{in} + 4(m - m_{in}) = |I| + 4m - m_{in}$. Since $C = k + 4m$ we get that $|I| - m_{in} \geq k$. Now we can transform $I$ into an independent set $I'$ by going over every one of the $m_{in}$ edges that are inside $I$ and removing one of the endpoints of the edge from it. The resulting set $I'$ is an independent set in the graph $G$ of size $|I| - m_{in} \geq k$ and so this concludes the proof of the soundness condition.

13.6 REDUCING 3SAT TO LONGEST PATH

One of the most basic algorithms in Computer Science is Dijkstra’s algorithm to find the shortest path between two vertices. We now show that in contrast, an efficient algorithm for the longest path problem would imply a polynomial-time algorithm for 3SAT.

**Theorem 13.7 — Hardness of longest path.**

$$3SAT \leq_p LONGPATH$$ (13.6)

**Proof Idea:**

To prove Theorem 13.7 need to show how to transform a 3CNF formula $\varphi$ into a graph $G$ and two vertices $s, t$ such that $G$ has a path of length at least $k$ if and only if $\varphi$ is satisfiable. The idea of the reduction is sketched in Fig. 13.7 and Fig. 13.8. We will construct a graph that contains a potentially long “snaking path” that corresponds to all variables in the formula. We will add a “gadget” corresponding to each clause of $\varphi$ in a way that we would only be able to use the gadgets if we have a satisfying assignment.
*Proof of Theorem 13.7.* We build a graph $G$ that “snakes” from $s$ to $t$ as follows. After $s$ we add a sequence of $n$ long loops. Each loop has an “upper path” and a “lower path”. A simple path cannot take both the upper path and the lower path, and so it will need to take exactly one of them to reach $s$ from $t$.

Our intention is that a path in the graph will correspond to an assignment $x \in \{0, 1\}^n$ in the sense that taking the upper path in the $i^{th}$ loop corresponds to assigning $x_i = 1$ and taking the lower path corresponds to assigning $x_i = 0$. When we are done snaking through all the $n$ loops corresponding to the variables to reach $t$ we need to pass through $m$ “obstacles”: for each clause $j$ we will have a small gadget consisting of a pair of vertices $s_j, t_j$ that have three paths between them. For example, if the $j^{th}$ clause had the form $x_{17} \lor x_{55} \lor x_{72}$ then one path would go through a vertex in the lower loop corresponding to $x_{17}$, one path would go through a vertex in the upper loop corresponding to $x_{55}$ and the third would go through the lower loop corresponding to $x_{72}$. We see that if we went in the first stage according to a satisfying assignment then we will be able to find a free vertex to travel from $s_j$ to $t_j$. We link $t_1$ to $s_2$, $t_2$ to $s_3$, etc and link $t_m$ to $t$. Thus a satisfying assignment would correspond to a path from $s$ to $t$ that goes through one path in each loop corresponding to the variables, and one path in each loop corresponding to the clauses. We can make the loop corresponding to the variables long enough so that we must take the entire path in each loop in order to have a fighting chance of getting a path as long as the one corresponds to a satisfying assignment. But if we do that, then the only way if we are able to reach $t$ is if the paths we took corresponded to a satisfying assignment, since otherwise we will have one clause $j$ where we cannot reach $t_j$ from $s_j$ without using a vertex we already used before.

\[\]

**Lecture Recap**

- The computational complexity of many seemingly unrelated computational problems can be related to one another through the use of reductions.
- If $F \leq_p G$ then a polynomial-time algorithm for $G$ can be transformed into a polynomial-time algorithm for $F$.
- Equivalently, if $F \leq_p G$ and $F$ does not have a polynomial-time algorithm then neither does $G$.
- We’ve developed many techniques to show that $3SAT \leq_p F$ for interesting functions $F$. Sometimes

---

*Figure 13.7: We can transform a 3SAT formula $\varphi$ into a graph $G$ such that the longest path in the graph $G$ would correspond to a satisfying assignment in $\varphi$. In this graph, the black colored part corresponds to the variables of $\varphi$ and the blue colored part corresponds to the vertices. A sufficiently long path would have to first “snake” through the black part, for each variable choosing either the “upper path” (corresponding to assigning it the value $True$) or the “lower path” (corresponding to assigning it the value $False$). Then to achieve maximum length the path would traverse through the blue part, where to go between two vertices corresponding to a clause such as $x_{17} \lor x_{32} \lor x_{57}$, the corresponding vertices would have to have been not traversed before.*

*Figure 13.8: The graph above with the longest path marked on it, the part of the path corresponding to variables is in green and part corresponding to the clauses is in pink.*
we can do so by using transitivity of reductions: if 3SAT \leq_p G and G \leq_p F then 3SAT \leq_p F.

13.7 EXERCISES

Exercise 13.1 — Transitivity of reductions. Prove that if F \leq_p G and G \leq_p H then F \leq_p H.

13.8 BIBLIOGRAPHICAL NOTES

Several notions of reductions are defined in the literature. The notion defined in Definition 13.1 is often known as a mapping reduction, many to one reduction or a Karp reduction.

The maximal (as opposed to maximum) independent set is the task of finding a “local maximum” of an independent set: an independent set S such that one cannot add a vertex to it without losing the independence property (such a set is known as a vertex cover). Unlike finding a maximum independent set, finding a maximal independent set can be done efficiently by a greedy algorithm, but this local maximum can be much smaller than the global maximum.

Reduction of independent set to max cut taken from these notes. Image of Hamiltonian Path through Dodecahedron by Christoph Sommer.

We have mentioned that the line between reductions used for algorithm design and showing hardness is sometimes blurry. An excellent example for this is the area of SAT Solvers (see [Gom+08]). In this field people use algorithms for SAT (that take exponential time in the worst case but often are much faster on many instances in practice) together with reductions of the form F \leq_p SAT to derive algorithms for other functions F of interest.

\textsuperscript{4} TODO: Maybe mention either in exercise or in body of the lecture some NP hard results motivated by science. For example, shortest superstring that is motivated by genome sequencing, protein folding, maybe others.
14

NP, NP completeness, and the Cook-Levin Theorem

“In this paper we give theorems that suggest, but do not imply, that these problems, as well as many others, will remain intractable perpetually”, Richard Karp, 1972

“Sad to say, but it will be many more years, if ever before we really understand the Mystical Power of Twoness... 2-SAT is easy, 3-SAT is hard, 2-dimensional matching is easy, 3-dimensional matching is hard. Why? oh, Why?” Eugene Lawler

14.1 THE CLASS NP

So far we have shown that 3SAT is no harder than Quadratic Equations, Independent Set, Maximum Cut, and Longest Path. But to show that these problems are computationally equivalent we need to give reductions in the other direction, reducing each one of these problems to 3SAT as well. It turns out we can reduce all three problems to 3SAT in one fell swoop.

In fact, this result extends far beyond these particular problems. All of the problems we discussed in Chapter 13, and a great many other problems, share the same commonality: they are all search problems, where the goal is to decide, given an instance $x$, whether there exists a solution $y$ that satisfies some condition that can be verified in polynomial time. For example, in 3SAT, the instance is a formula and the solution is an assignment to the variable; in Max-Cut the instance is a graph and the solution is a cut in the graph; and so on and so forth. It turns out that every such search problem can be reduced to 3SAT.

To make this precise, we make the following mathematical definition: we define the class NP to contain all Boolean functions that correspond to a search problem of the form above— that is, functions that output 1 on $x$ if and only if there exists a solution $w$ such that
the pair \((x, w)\) satisfies some polynomial-time checkable condition. Formally, \(\text{NP}\) is defined as follows:

**Definition 14.1 — NP.** We say that \(F : \{0, 1\}^* \rightarrow \{0, 1\}\) is in \(\text{NP}\) if there exists some constants \(a, b \in \mathbb{N}\) and \(V : \{0, 1\}^* \rightarrow \{0, 1\}\) such that \(V \in \text{P}\) and for every \(x \in \{0, 1\}^n\),

\[
F(x) = 1 \iff \exists w \in \{0, 1\}^a \text{ s.t. } V(xw) = 1. \quad (14.1)
\]

In other words, for \(F\) to be in \(\text{NP}\), there needs to exist some polynomial-time computable verification function \(V\), such that if \(F(x) = 1\) then there must exist \(w\) (of length polynomial in \(|x|\)) such that \(V(xw) = 1\), and if \(F(x) = 0\) then for every such \(w\), \(V(xw) = 0\). Since the existence of this string \(w\) certifies that \(F(x) = 1\), \(w\) is often referred to as a certificate, witness, or proof that \(F(x) = 1\).

See also Fig. 14.1 for an illustration of Definition 14.1. The name \(\text{NP}\) stands for “nondeterministic polynomial time” and is used for historical reasons; see the bibliographical notes. The string \(w\) in (14.1) is sometimes known as a solution, certificate, or witness for the instance \(x\).

**Remark 14.2 — NP and proof systems.** The definition of \(\text{NP}\) means that for every \(F \in \text{NP}\) and string \(x \in \{0, 1\}^*\), \(F(x) = 1\) if and only if there is a short and efficiently verifiable proof of this fact. That is, we can think of the function \(V\) in Definition 14.1 as a verifier algorithm, similar to what we’ve seen in Section 10.1. The verifier checks whether a given string \(w \in \{0, 1\}^*\) is a valid proof for the statement “\(F(x) = 1\)”. Essentially all proof systems considered in mathematics involve line-by-line checks that can be carried out in polynomial time. Thus the heart of \(\text{NP}\) is asking for statements that have short (i.e., polynomial in the size of the statements) proof. Indeed, as we will see in Chapter 15, Kurt Gödel phrased the question of whether \(\text{NP} = \text{P}\) as asking whether “the mental work of a mathematician [in proving theorems] could be completely replaced by a machine”.

**Definition 14.1** is asymmetric in the sense that there is a difference between an output of 1 and an output of 0. You should make sure you understand why this definition does not guarantee that if \(F \in \text{NP}\) then the function \(1 - F\) (i.e., the map \(x \mapsto 1 - F(x)\)) is in \(\text{NP}\) as well. In fact, it is believed that there do exist functions such that \(1 - F \notin \text{NP}\) even when \(F \in \text{NP}\).
14.1.1 Examples of NP functions

Example 14.3 — 3SAT ∈ NP. 3SAT is in NP since for every ℓ-variable formula \( \varphi \), \( 3\text{SAT}(\varphi) = 1 \) if and only if there exists a satisfying assignment \( x \in \{0, 1\}^\ell \) such that \( \varphi(x) = 1 \), and we can check this condition in polynomial time.

The above reasoning explains why 3SAT is in NP, but since this is our first example, we will now belabor the point and expand out in full formality what is the precise representation of the witness \( w \) and the algorithm \( V \) that demonstrate that 3SAT is in NP.

Specifically, we can represent a 3CNF formula \( \varphi \) with \( k \) variables and \( m \) clauses as a string of length \( n = O(m \log k) \), since every one of the \( m \) clauses involves three variables and their negation, and the identity of each variable can be represented using \( \lceil \log_2 k \rceil \). We assume that every variable participates in some clause (as otherwise it can be ignored) and hence that \( m \geq k \), which in particular means that \( n \) is larger than both \( m \) and \( k \).

We can represent an assignment to the \( k \) variables using a \( k \)-length string, which, since \( n > k \), can be “padded” to a string \( w \in \{0, 1\}^n \) in some standard way. (For example, if \( y \in \{0, 1\}^k \) is the assignment, we can let \( w = y10^{n-k-1} \); given the string \( w \) we can “read off” \( y \), by chopping off all the zeroes at the end of \( w \) until we encounter the first 1, which we remove as well.)

Now checking whether a given assignment \( y \in \{0, 1\}^k \) satisfies a given \( k \)-variable 3CNF \( \varphi \) can be done in polynomial time through the following algorithm \( V \):

---

1 For example, as shown below, 3SAT ∈ NP, but the function 3SAT that on input a 3CNF formula \( \varphi \) outputs 1 if and only if \( \varphi \) is not satisfiable is not known (nor believed) to be in NP.
Algorithm $V$:

**Input:**

1. 3CNF formula $\varphi$ with $k$ variables and $m$ clauses (encoded as a string of length $n = O(m \log k)$)
2. Assignment $y \in \{0, 1\}^k$ to the variables of $\varphi$ (encoded using padding as a string $w \in \{0, 1\}^n$)

**Output:** 1 if and only if $y$ satisfies $\varphi$.

**Operation:**

1. For every clause $C = (\ell_1 \lor \ell_2 \lor \ell_3)$ of $\varphi$ (where $\ell_1, \ell_2, \ell_3$ are literals), if all three literals evaluate to false under the assignment $y$ then halt and output 0.
2. Output 1.

Algorithm $V$ runs in time polynomial in the length $n$ of $\varphi$’s description as a string. Indeed there are $m$ clauses, and checking the evaluation of a literal of the form $y_i$ or $\neg y_j$ can be done by scanning the $k$-length string $y$, and hence the running time of Algorithm $V$ is at most $O(mk) = O(n^2)$, as both $k$ and $m$ are smaller than $n$.

By its definition the algorithm outputs 1 if and only if the assignment $y$ satisfies all the clauses of the 3CNF formula $\varphi$, which means that $3SAT(\varphi) = 1$ if and only if there exists some $w \in \{0, 1\}^n$ such that $V(\varphi w) = 1$ which is precisely the condition needed to show that $3SAT \in \text{NP}$ per Definition 14.1.

**Remark 14.4 — Padding a witness.** The “padding trick” we used in Example 14.3 can always be used to expand a witness of length smaller than $an^b$ to a witness of exactly that length. Therefore one can think of the condition (14.1) in Definition 14.1 as simply stipulating that the “solution” $w$ to the problem $x$ is of length at most polynomial in $|x|$.

Here are some more examples for problems in NP. For each one of these problems we merely sketch how the witness is represented and why it is efficiently checkable, but working out the details can be a good way to get more comfortable with Definition 14.1:
• **QUADEQ** is in \textbf{NP} since for every \(\ell\)-variable instance of quadratic equations \(E\), \(\text{QUADEQ}(E) = 1\) if and only if there exists an assignment \(x \in \{0, 1\}^\ell\) that satisfies \(E\). We can check the condition that \(x\) satisfies \(E\) in polynomial time by enumerating over all the equations in \(E\), and for each such equation \(e\), plug in the values of \(x\) and verify that \(e\) is satisfied.

• **ISET** is in \textbf{NP} since for every graph \(G\) and integer \(k\), \(\text{ISET}(G, k) = 1\) if and only if there exists a set \(S\) of \(k\) vertices that contains no pair of neighbors in \(G\). We can check the condition that \(S\) is an independent set of size \(\geq k\) in polynomial time by first checking that \(|S| \geq k\) and then enumerating over all edges \(\{u, v\}\) in \(G\), and for each such edge verify that either \(u \notin S\) or \(v \notin S\).

• **LONGPATH** is in \textbf{NP} since for every graph \(G\) and integer \(k\), \(\text{LONGPATH}(G, k) = 1\) if and only if there exists a simple path \(P\) in \(G\) that is of length at least \(k\). We can check the condition that \(P\) is a simple path of length \(k\) in polynomial time by checking that it has the form \((v_0, v_1, \ldots, v_k)\) where each \(v_i\) is a vertex in \(G\), no \(v_i\) is repeated, and for every \(i \in [k]\), the edge \(\{v_i, v_{i+1}\}\) is present in the graph.

• **MAXCUT** is in \textbf{NP} since for every graph \(G\) and integer \(k\), \(\text{MAXCUT}(G, k) = 1\) if and only if there exists a cut \((S, \overline{S})\) in \(G\) that cuts at least \(k\) edges. We can check that condition that \((S, \overline{S})\) is a cut of value at least \(k\) in polynomial time by checking that \(S\) is a subset of \(G\)'s vertices and enumerating over all the edges \(\{u, v\}\) of \(G\), counting those edges such that \(u \in S\) and \(v \notin S\) or vice versa.

14.1.2 Basic facts about \textbf{NP}

The definition of \textbf{NP} is one of the most important definitions of this book, and is worth while taking the time to digest and internalize. The following solved exercises establish some basic properties of this class. As usual, I highly recommend that you try to work out the solutions yourself.

\textbf{Solved Exercise 14.1} — Verifying is no harder than solving. Prove that \(\mathbb{P} \subseteq \textbf{NP}\).

\textbf{Solution:}

Suppose that \(F \in \mathbb{P}\). Define the following function \(V: V(x0^n) = 1\) iff \(n = |x|\) and \(F(x) = 1\). \((V\) outputs 0 on all other inputs.) Since \(F \in \mathbb{P}\) we can clearly compute \(V\) in polynomial time as well.

Let \(x \in \{0, 1\}^n\) be some string. If \(F(x) = 1\) then \(V(x0^n) = 1\). On the other hand, if \(F(x) = 0\) then for every \(w \in \{0, 1\}^n, V(xw) = 0\).
Therefore, setting \( a = b = 1 \), we see that \( V \) satisfies (14.1), and establishes that \( F \in \textbf{NP} \).

**Remark 14.5 — \( \textbf{NP} \) does not mean non-polynomial!!**

People sometimes think that \( \textbf{NP} \) stands for “non-polynomial time”. As Solved Exercise 14.1 shows, this is far from the truth, and in fact every polynomial-time computable function is in \( \textbf{NP} \) as well.

If \( F \) is in \( \textbf{NP} \) it certainly does not mean that \( F \) is hard to compute (though it does not, as far as we know, necessarily mean that it’s easy to compute either). Rather, it means that \( F \) is easy to verify, in the technical sense of Definition 14.1.

**Solved Exercise 14.2 — \( \textbf{NP} \) is in exponential time.** Prove that \( \textbf{NP} \subseteq \textbf{EXP} \).

**Solution:**

Suppose \( F \in \textbf{NP} \) and let \( V \) be the polynomial-time computable function that satisfies (14.1) and \( a, b \) the corresponding constants. Then the following is an exponential-time algorithm \( A \) to compute \( F \):

**Algorithm \( A \):**

**Input:** \( x \in \{0, 1\}^* \), let \( n = |x| \)

**Operation:**

1. For every \( w \in \{0, 1\}^{anb} \), if \( V(xw) = 1 \) then halt and output 1.
2. Output 0.

Since \( V \in \textbf{P} \), for every \( x \in \{0, 1\}^n \), Algorithm \( A \) runs in time \( \text{poly}(n)2^{anb} \). Moreover by (14.1), \( A \) will output 1 on \( x \) if and only if \( F(x) = 1 \).

Solved Exercise 14.1 and Solved Exercise 14.2 together imply that

\[
\textbf{P} \subseteq \textbf{NP} \subseteq \textbf{EXP}.
\]  

(14.2)

The time hierarchy theorem (Theorem 12.8) implies that \( \textbf{P} \subseteq \textbf{EXP} \) and hence at least one of the two inclusions \( \textbf{P} \subseteq \textbf{NP} \) or \( \textbf{NP} \subseteq \textbf{EXP} \) is strict. It is believed that both of them are in fact strict inclusions. That is, it is believed that there are functions in \( \textbf{NP} \) that cannot be
computed in polynomial time (this is the \( P \neq NP \) conjecture) and that there are functions \( F \) in \( EXP \) for which we cannot even efficiently certify that \( F(x) = 1 \) for a given input \( x \).\(^2\)

We have previously informally equated the notion of \( F \leq_p G \) with \( F \) being “no harder than \( G \” \) and in particular have seen in Solved Exercise 13.1 that if \( G \in P \) and \( F \leq_p G \), then \( F \in P \) as well. The following exercise shows that if \( F \leq_p G \) then it is also “no harder to verify” than \( G \). That is, regardless of whether or not it is in \( P \), if \( G \) has the property that solutions to it can be efficiently verified, then so does \( F \).

**Solved Exercise 14.3 — Reductions and \( NP \).** Let \( F, G : \{0, 1\}^* \rightarrow \{0, 1\} \).

Show that if \( F \leq_p G \) and \( G \in NP \) then \( F \in NP \).

**Solution:**

Suppose that \( G \) is in \( NP \) and in particular there exists \( a, b \) and \( V \in P \) such that for every \( y \in \{0, 1\}^* \), \( G(y) = 1 \iff \exists w \in \{0, 1\}^a|y|^b V(yw) = 1 \). Define \( V'(x, w) = 1 \) iff \( V(R(x)w) = 1 \) where \( R \) is the polynomial-time reduction demonstrating that \( F \leq_p G \). Then for every \( x \in \{0, 1\}^* \),

\[
F(x) = 1 \iff G(R(x)) = 1 \iff \exists w \in \{0, 1\}^a|R(x)|^b V(R(x)w) = 1 \iff \exists w \in \{0, 1\}^a|R(x)|^b V'(x, w) = 1
\]

(14.3)

Since there are some constants \( a', b' \) such that \( |R(x)| \leq a'|x|^{b'} \) for every \( x \in \{0, 1\}^* \), by simple padding we can modify \( V' \) to an algorithm that certifies that \( F \in NP \).\(^2\)

14.2 FROM \( NP \) TO 3SAT: THE COOK-LEVIN THEOREM

We have seen everal example of problems for which we do not know if their best algorithm is polynomial or exponential, but we can show that they are in \( NP \). That is, we don’t know if they are easy to solve, but we do know that it is easy to verify a given solution. There are many, many, many, more examples of interesting functions we would like to compute that are easily shown to be in \( NP \). What is quite amazing is that if we can solve 3SAT then we can solve all of them!

The following is one of the most fundamental theorems in Computer Science:

**Theorem 14.6 — Cook-Levin Theorem.** For every \( F \in NP \), \( F \leq_p 3SAT \).

We will soon show the proof of Theorem 14.6, but note that it immediately implies that \( QUADEQ, LONGPATH \), and \( MAXCUT \) all

\(^2\) One function \( F \) that is believed to lie outside \( NP \) is the function \( 3SAT \) defined as \( 3SAT(\varphi) = 1 - 3SAT(\varphi) \) for every 3CNF formula \( \varphi \). The conjecture that \( 3SAT \notin NP \) is known as the “\( NP \neq coNP \)” conjecture. It implies the \( P \neq NP \) conjecture (can you see why?).
reduce to 3SAT. Combining it with the reductions we’ve seen in Chapter 13, it implies that all these problems are equivalent! For example, to reduce QUADEQ to LONGPATH, we can first reduce QUADEQ to 3SAT using Theorem 14.6 and use the reduction we’ve seen in Theorem 13.7 from 3SAT to LONGPATH. That is, since QUADEQ ∈ NP, Theorem 14.6 implies that QUADEQ ≤_p 3SAT, and Theorem 13.7 implies that 3SAT ≤_p LONGPATH, which by the transitivity of reductions (Lemma 13.2) means that QUADEQ ≤_p LONGPATH. Similarly, since LONGPATH ∈ NP, we can use Theorem 14.6 and Theorem 13.4 to show that LONGPATH ≤_p 3SAT ≤_p QUADEQ, concluding that LONGPATH and QUADEQ are computationally equivalent.

There is of course nothing special about QUADEQ and LONGPATH here: by combining (14.6) with the reductions we saw, we see that just like 3SAT, every F ∈ NP reduces to LONGPATH, and the same is true for QUADEQ and MAXCUT. All these problems are in some sense “the hardest in NP” since an efficient algorithm for any one of them would imply an efficient algorithm for all the problems in NP. This motivates the following definition:

**Definition 14.7 — NP-hardness and NP-completeness.** We say that \( G : \{0, 1\}^* \rightarrow \{0, 1\} \) is NP hard if for every \( F \in \text{NP} \), \( F \leq_p G \).

We say that \( G : \{0, 1\}^* \rightarrow \{0, 1\} \) is NP complete if \( G \) is NP hard and \( G \) is in NP.

The Cook-Levin Theorem (Theorem 14.6) can be rephrased as saying that 3SAT is NP hard, and since it is also in NP, this means that 3SAT is NP complete. Together with the reductions of Chapter 13, Theorem 14.6 shows that despite their superficial differences, 3SAT, quadratic equations, longest path, independent set, and maximum cut, are all NP-complete. Many thousands of additional problems have been shown to be NP-complete, arising from all the sciences, mathematics, economics, engineering and many other fields.

14.2.1 What does this mean?
As we’ve seen in Solved Exercise 14.1, \( P \subseteq \text{NP} \). The most famous conjecture in Computer Science is that this containment is strict. That is, it is widely conjectured that \( P \neq \text{NP} \). One way to refute the conjecture that \( P \neq \text{NP} \) is to give a polynomial-time algorithm for even a single one of the NP-complete problems such as 3SAT, Max Cut, or the thousands of others that have been studied in all fields of human endeavors. The fact that these problems have been studied by so many people, and yet not a single polynomial-time algorithm for any of them has been found, supports that conjecture that indeed \( P \neq \text{NP} \). In fact, for many of these problems (including all the ones we mentioned above), we don’t even know of a \( 2^{o(n)} \)-time algorithm! However, to the
frustration of computer scientists, we have not yet been able to prove that \( P \neq NP \) or even rule out the existence of an \( O(n) \)-time algorithm for 3SAT. Resolving whether or not \( P = NP \) is known as the P vs NP problem. A million-dollar prize has been offered for the solution of this problem, a popular book has been written, and every year a new paper comes out claiming a proof of \( P = NP \) or \( P \neq NP \), only to wither under scrutiny.\(^4\) The following 120 page survey of Aaronson, as well as chapter 3 in Wigderson’s upcoming book are excellent sources for summarizing what is known about this problem.

One of the mysteries of computation is that people have observed a certain empirical “zero-one law” or “dichotomy” in the computational complexity of natural problems, in the sense that many natural problems are either in \( P \) (often in \( TIME(O(n)) \) or \( TIME(O(n^2)) \)), or they are are \( NP \) hard. This is related to the fact that for most natural problems, the best known algorithm is either exponential or polynomial, with not too many examples where the best running time is some strange intermediate complexity such as \( 2^{\sqrt{n}} \). However, it is believed that there exist problems in \( NP \) that are neither in \( P \) nor are \( NP \)-complete, and in fact a result known as “Ladner’s Theorem” shows that if \( P \neq NP \) then this is indeed the case (see also Exercise 14.1 and Fig. 14.2).

### 14.2.2 The Cook-Levin Theorem: Proof outline

We will now prove the Cook-Levin Theorem, which is the underpinning to a great web of reductions from 3SAT to thousands of problems across great many fields. Some problems that have been shown to be \( NP \)-complete include: minimum-energy protein folding, minimum surface-area foam configuration, map coloring, optimal Nash equilibrium, quantum state entanglement, minimum supersequence of a genome, minimum codeword problem, shortest vector in a lattice, minimum genus knots, positive Diophantine equations, integer programming, and many many more. The worst-case complexity of all these problems is (up to polynomial factors) equivalent to that of 3SAT, and through the Cook-Levin Theorem, to all problems in \( NP \).

To prove Theorem 14.6 we need to show that \( P \leq_p 3SAT \) for every \( F \in NP \). We will do so in three stages. We define two intermediate problems: NANDSAT and 3NAND. We will shortly show the definitions of these two problems, but Theorem 14.6 will follow from combining the following three results:

1. **NANDSAT is NP hard** (Lemma 14.8).
2. **NANDSAT \( \leq_p 3NAND \)** (Lemma 14.10).

\(^4\) The following web page keeps a catalog of these failed attempts. At the time of this writing, it lists about 110 papers claiming to resolve the question, of which about 60 claim to prove that \( P = NP \) and about 50 claim to prove that \( P \neq NP \).

\(^5\) TODO: maybe add examples of NP hard problems as a barrier to understanding - problems from economics, physics, etc.. that prevent having a closed-form solutions

\(^6\) TODO: maybe include knots
3. $3\text{NAND} \leq_p 3\text{SAT}$ (Lemma 14.11).

By the transitivity of reductions, it will follow that for every $F \in \text{NP}$,

$$F \leq_p \text{NANDSAT} \leq_p 3\text{NAND} \leq_p 3\text{SAT}$$

hence establishing Theorem 14.6.

We will prove these three results Lemma 14.8, Lemma 14.10 and Lemma 14.11 one by one, providing the requisite definitions as we go along.

### 14.3 THE NANDSAT PROBLEM, AND WHY IT IS NP HARD.

We define the $\text{NANDSAT}$ problem as follows. On input a string $Q \in \{0, 1\}^*$, we define $\text{NANDSAT}(Q) = 1$ if and only if $Q$ is a valid representation of an $n$-input and single-output NAND program and there exists some $w \in \{0, 1\}^n$ such that $Q(w) = 1$. While we don’t need this to prove Lemma 14.8, note that $\text{NANDSAT}$ is in $\text{NP}$ since we can verify that $Q(w) = 1$ using the polynomial-time algorithm for evaluating NAND-CIRC programs.\(^7\) We now prove that $\text{NANDSAT}$ is NP hard.

**Lemma 14.8** $\text{NANDSAT}$ is NP hard.

**Proof Idea:**

To prove Lemma 14.8 we need to show that for every $F \in \text{NP}$, $F \leq_p \text{NANDSAT}$. The high-level idea is that by the definition of $\text{NP}$, there is some NAND-TM program $P^*$ and some polynomial $T(\cdot)$ such that $F(x) = 1$ if and only if there exists some $w \in \{0, 1\}^{ax+b}$ such that $P^*(xw)$ outputs 1 within $T(|x|)$ steps. Now by “unrolling the loop” of the NAND-TM program $P^*$ we can convert it into an $O(T(n))$ NAND-CIRC program $Q'$ with $n + an^b$ inputs and a single output such that for every $x \in \{0, 1\}^n$ and $w \in \{0, 1\}^{an^b}$, $Q'(xw) = P^*(xw)$. on input $x \in \{0, 1\}^n$ that on input $w$ will simulate $P^*(xw)$ for $T(|x|)$ steps. The next step is to hardwire the input $x$ to $Q'$ to obtain an $O(T(n))$ line NAND-CIRC program $Q$ with $m = an^b$ inputs such that for every $w \in \{0, 1\}^m$, $Q'(w) = Q(xw)$. By construction it will be the case that for every $x \in \{0, 1\}^n$, $F(x) = 1$ if and only if there exists $w \in \{0, 1\}^{an^b}$ such that $Q(w) = 1$, and hence this shows that $F \leq_p \text{NANDSAT}$.

\(^7\) $Q$ is a NAND-CIRC program and not a NAND-TM program, and hence it is only defined on inputs of some particular size $n$. Evaluating $Q$ on any input $w \in \{0, 1\}^n$ can be done in time polynomial in the number of lines of $Q$. 

---

The proof is a little bit technical but ultimately follows quite directly from the definition of $\text{NP}$, as well as of NAND and NAND-TM programs. If you find it confusing, try to pause here and work out the proof.
you yourself from these definitions, using the idea of “unravelling the loop” of a NAND-TM program. It might also be useful for you to think how you would implement in your favorite programming language the function unroll which on input a NAND-TM program \( P \) and numbers \( T, n \) would output an \( n \)-input NAND-CIRC program \( Q \) of \( O(|T|) \) lines such that for every input \( z \in \{0, 1\}^n \), if \( P \) halts on \( z \) within at most \( T \) steps and outputs \( y \), then \( Q(z) = y \).

**Proof of Lemma 14.8.** We now present the details. Let \( F \in \text{NP} \). To prove Lemma 14.8 we need to give a polynomial-time computable function that will map every \( x^* \in \{0, 1\}^n \) to a NAND-CIRC program \( Q \) such that \( F(x) = \text{NANDSAT}(Q) \).

Let \( x^* \in \{0, 1\}^n \) be such a string and let \( n = |x^*| \) be its length. By Definition 14.1 there exists \( V \in \text{P} \) and \( a, b \in \mathbb{N} \) such that \( F(x^*) = 1 \) if and only if there exists \( w \in \{0, 1\}^m \) such that \( V(x^*w) = 1 \).

Let \( m = an^b \). Since \( V \in \text{P} \) there is some NAND-TM program \( P^* \) that computes \( V \) on inputs of the form \( xw \) with \( x \in \{0, 1\}^n \) and \( w \in \{0, 1\}^m \) in at most \( (n + m)^c \) time for some constant \( c \). Using our “unrolling the loop NAND-TM to NAND compiler” of Theorem 12.12, we can obtain a NAND-CIRC program \( Q^* \) that has \( n + m \) inputs and at most \( O((n + m)^c) \) lines such that \( Q^*(xw) = P^*(xw) \) for every \( x \in \{0, 1\}^n \) and \( w \in \{0, 1\}^m \).

Now we can use the following simple but useful “hardwiring” technique to obtain a program:

**Lemma 14.9 — Hardwiring Lemma.** Given a \( T \)-line NAND-CIRC program \( Q^\prime \) of \( n + m \) inputs and \( x^* \in \{0, 1\}^n \), we can obtain in polynomial a program \( Q \) with \( m \) inputs of \( T + 3 \) lines such that for every \( w \in \{0, 1\}^m \), \( Q(w) = Q^\prime(x^*w) \).

**Proof.** To compute \( Q \), we simply do a “search and replace” for all references in \( Q^\prime \) to \( X[i] \) for \( i \in [n] \), and transform them to either the variable \( 0 \) or \( 1 \) depending on whether \( x_i^* \) is equal to 0 or 1 respectively. By adding three lines to the beginning of \( Q^\prime \), we can ensure that the \( 0 \) and \( 1 \) variables will have the correct value. The only thing that then remains to do another search and replace to transform all references to the variables \( X[n], \ldots, X[n + m - 1] \) to the variables \( X[0], \ldots, X[m - 1] \) so that the \( m \) inputs to the new program \( Q \) will correspond to last \( m \) inputs of the original program \( Q^\prime \). See Fig. 14.4 for an implementation of this reduction in Python.

Using Lemma 14.9, we obtain a program \( Q \) of \( m \) inputs such that \( Q(w) = Q^\prime(x^*w) = P^*(x^*w) \) for every \( w \in \{0, 1\}^m \). Since we know that
\( F(x^*) = 1 \) if and only if there exists \( w \in \{0, 1\}^m \) such that \( P^*(x^*w) = 1 \), this means that \( F(x^*) = 1 \) if and only if \( \text{NANDSAT}(Q) = 1 \), which is what we wanted to prove.

### 14.4 THE 3NAND PROBLEM

The 3NAND problem is defined as follows: the input is a logical formula \( \varphi \) on a set of variables \( z_0, \ldots, z_{r-1} \) which is an AND of constraints of the form \( z_i = \text{NAND}(z_j, z_k) \). For example, the following is a 3NAND formula with 5 variables and 3 constraints:

\[
(z_3 = \text{NAND}(z_0, z_2)) \land (z_1 = \text{NAND}(z_0, z_2)) \land (z_4 = \text{NAND}(z_3, z_1))
\]  

(14.5)

The output of 3NAND on input \( \varphi \) is 1 if and only if there is an assignment to the variables of \( \varphi \) that makes it evaluate to “true” (that is, there is some assignment \( z \in \{0, 1\}^r \) satisfying all of the constraints of \( \varphi \)). As usual, we can represent \( \varphi \) as a string, and so think of 3NAND as a function mapping \( \{0, 1\}^* \) to \( \{0, 1\} \). We now prove that 3NAND is NP hard:

**Lemma 14.10** \( \text{NANDSAT} \leq_p \text{3NAND} \).

**Proof Idea:**

To prove Lemma 14.10 we need to give a polynomial-time map from every NAND-CIRC program \( Q \) to a 3NAND formula \( \Psi \) such that there exists \( w \) such that \( Q(w) = 1 \) if and only if there exists \( z \) satisfying \( \Psi \). For every line \( i \) of \( Q \), we define a corresponding variable \( z_i \) of \( \Psi \). If the line \( i \) has the form \( \text{foo} = \text{NAND}(\text{bar}, \text{blah}) \) then we will add the clause \( z_i = \text{NAND}(z_j, z_k) \) where \( j \) and \( k \) are the last lines in which \( \text{bar} \) and \( \text{blah} \) were written to. We will also set variables corresponding to the input variables, as well as add a clause to ensure that the final...
output is 1. The resulting reduction can be implemented in about a dozen lines of Python, see ??.

```python
# Reduce NAND to 3NAND
# Output: NAND formula \( \Psi \)
# Input: NAND-CIRC program \( Q \)
# Variables: \( z_0, \ldots, z_{m+n-1} \)

u = NAND(x[0], x[1])
v = NAND(x[0], u)
w = NAND(x[1], u)
x = NAND(x[1], v)
y = NAND(x[2], u)
z = NAND(v, w)
w0 = NAND(v, w)
x0 = NAND(v, w)
y0 = NAND(x[0], u)
z0 = NAND(x[0], u)
\( y[0] = \text{NAND}(v, w) \)

print(\text{NANDSAT\_to\_3NAND}(xor5))

(xor5 = NAND(x[0], x[1]) \land (x[0] = \text{NAND}(x[0], \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1]))) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])) \land (x[0] = \text{NAND}(x[0], x[1])))

eval(\text{NANDSAT\_to\_3NAND}(xor5))

True
```

**Figure 14.5:** Python code to reduce an instance \( Q \) of NANDSAT to an instance \( \Psi \) of 3NAND. In the example above we transform the NAND-CIRC program xor5 which has 5 input variables and 16 lines, into a 3NAND formula \( \Psi \) that has 24 variables and 20 clauses. Since xor5 outputs 1 on the input 1, 0, 0, 1, 1, there exists an assignment \( z \in \{0, 1\}^{24} \) to \( \Psi \) such that \( (z[0], z[1], z[2], z[3], z[4]) = (1, 0, 0, 1, 1) \) and \( \Psi \) evaluates to \text{true} on \( z \).

**Proof of Lemma 14.10.** To prove Lemma 14.10 we need to give a reduction from NANDSAT to 3NAND. Let \( Q \) be a NAND-CIRC program with \( n \) inputs, one output, and \( m \) lines. We can assume without loss of generality that \( Q \) contains the variables one and zero as usual.

We map \( Q \) to a 3NAND formula \( \Psi \) as follows:

1. \( \Psi \) has \( m + n \) variables \( z_0, \ldots, z_{m+n-1} \).
2. The first \( n \) variables \( z_0, \ldots, z_{n-1} \) will correspond to the inputs of \( Q \). The next \( m \) variables \( z_n, \ldots, z_{n+m-1} \) will correspond to the \( m \) lines of \( Q \).
3. For every \( \ell \in \{n, n+1, \ldots, n+m\} \), if the \( \ell \) \(- n\)-th line of the program \( Q \) is \( \text{foo} = \text{NAND}(\text{bar}, \text{blah}) \) then we add to \( \Psi \) the constraint \( \zeta_\ell = \text{NAND}(z_j, z_k) \) where \( j - n \) and \( k - n \) correspond to the last lines in which the variables \( \text{bar} \) and \( \text{blah} \) (respectively) were written to. If one or both of \( \text{bar} \) and \( \text{blah} \) was not written to before then we use \( z_\ell \) instead of the corresponding value \( z_j \) or \( z_k \) in the constraint, where \( \ell - n \) is the line in which \( \text{zero} \) is assigned a value. If one or both of \( \text{bar} \) and \( \text{blah} \) is an input variable \( X[i] \) then we use \( z_i \) in the constraint.
4. Let \( \ell^* \) be the last line in which the output \( y_0 \) is assigned a value. Then we add the constraint \( \zeta_{\ell^*} = \text{NAND}(z_{\ell^*}, z_0) \) where \( \ell^* - n \) is as above the last line in which \( \text{zero} \) is assigned a value. Note that this is effectively the constraint \( \zeta_{\ell^*} = \text{NAND}(0, 0) = 1 \).
To complete the proof we need to show that there exists \( w \in \{0, 1\}^n \) s.t. \( Q(w) = 1 \) if and only if there exists \( z \in \{0, 1\}^{n+m} \) that satisfies all constraints in \( \Psi \). We now show both sides of this equivalence.

**Part I: Completeness.** Suppose that there is \( w \in \{0, 1\}^n \) s.t. \( Q(w) = 1 \). Let \( z \in \{0, 1\}^{n+m} \) be defined as follows: for \( i \in [n] \), \( z_i = w_i \) and for \( i \in \{n, n + 1, \ldots, n + m\} \) \( z_i \) equals the value that is assigned in the \((i - n)\)-th line of \( Q \) when executed on \( w \). Then by construction \( z \) satisfies all of the constraints of \( \Psi \) (including the constraint that \( z_{n+m} = \text{NAND}(0, 0) = 1 \) since \( Q(w) = 1 \)).

**Part II: Soundness.** Suppose that there exists \( z \in \{0, 1\}^{n+m} \) satisfying \( \Psi \). Soundness will follow by showing that \( Q(z_0, \ldots, z_{n-1}) = 1 \) (and hence in particular there exists \( w \in \{0, 1\}^n \), namely \( w = z_0 \cdots z_{n-1} \), such that \( Q(w) = 1 \)). To do this we will prove the following claim \((\ast)\): for every \( \ell \in [m] \), \( z_{\ell+n} \) equals the value assigned in the \( \ell \)-th step of the execution of the program \( Q \) on \( z_0, \ldots, z_{n-1} \). Note that because \( z \) satisfies the constraints of \( \Psi \), \((\ast)\) is sufficient to prove the soundness condition since these constraints imply that the last value assigned to the variable \( y_0 \) in the execution of \( Q \) on \( z_0 \cdots z_{n-1} \) is equal to 1. To prove \((\ast)\) suppose, towards a contradiction, that it is false, and let \( \ell \) be the smallest number such that \( z_{\ell+n} \) is not equal to the value assigned in the \( \ell \)-th step of the execution of \( Q \) on \( z_0, \ldots, z_{n-1} \). But since \( z \) satisfies the constraints of \( \Psi \), we get that \( z_{\ell+n} = \text{NAND}(z_i, z_j) \) where (by the assumption above that \( \ell \) is smallest with this property) these values do correspond to the values last assigned to the variables on the righthand side of the assignment operator in the \( \ell \)-th line of the program. But this means that the value assigned in the \( \ell \)-th step is indeed simply the \( \text{NAND} \) of \( z_i \) and \( z_j \), contradicting our assumption on the choice of \( \ell \).


14.5 **FROM 3NAND TO 3SAT**

To conclude the proof of Theorem 14.6, we need to show Lemma 14.11 and show that \( 3\text{NAND} \leq_p 3\text{SAT} \):

**Lemma 14.11** \( 3\text{NAND} \leq_p 3\text{SAT} \).

**Proof Idea:**

To prove Lemma 14.11 we need to map a 3NAND formula \( \varphi \) into a 3SAT formula \( \psi \) such that \( \varphi \) is satisfiable if and only if \( \psi \) is. The idea is that we can transform every NAND constraint of the form \( a = \text{NAND}(b, c) \) into the AND of ORs involving the variables \( a, b, c \) and their negations, where each of the ORs contains at most three terms. The construction is fairly straightforward, and the details are given below.
It is a good exercise for you to try to find a 3CNF formula $\xi$ on three variables $a, b, c$ such that $\xi(a, b, c)$ is true if and only if $a = NAND(b, c)$. Once you do so, try to see why this implies a reduction from 3NAND to 3SAT, and hence completes the proof of Lemma 14.11.

\[ z_i = NAND(z_j, z_k) \]  

(14.6)

is satisfied if $z_i = 1$ whenever $(z_j, z_k) \neq (1, 1)$. By going through all cases, we can verify that (14.6) is equivalent to the constraint

\[ (\overline{z_j} \lor \overline{z_j} \lor z_k) \land (z_i \lor z_j) \land (z_i \lor z_k) . \]  

(14.7)

Indeed if $z_j = z_k = 1$ then the first constraint of Eq. (14.7) is only true if $z_i = 0$. On the other hand, if either of $z_j$ or $z_k$ equals 0 then unless $z_i = 1$ either the second or third constraints will fail. This means that, given any 3NAND formula $\varphi$ over $n$ variables $z_0, \ldots, z_{n-1}$, we can obtain a 3SAT formula $\psi$ over the same variables by replacing every 3NAND constraint of $\varphi$ with three 3OR constraints as in Eq. (14.7). Because of the equivalence of (14.6) and (14.7), the formula $\psi$ satisfies that $\psi(z_0, \ldots, z_{n-1}) = \varphi(z_0, \ldots, z_{n-1})$ for every assignment $z_0, \ldots, z_{n-1} \in \{0, 1\}^n$ to the variables. In particular $\psi$ is satisfiable if and only if $\varphi$ is, thus completing the proof.

14.6 WRAPPING UP

We have shown that for every function $F$ in $\textbf{NP}$, $F \leq_p \textbf{NANDSAT}$, $3\text{NAND} \leq_p 3\text{SAT}$, and so $3\text{SAT}$ is $\textbf{NP}$-hard. Since in Chapter 13 we saw that $3\text{SAT} \leq_p \textbf{QUADEQ}$, $3\text{SAT} \leq_p \text{ISET}$, $3\text{SAT} \leq_p \text{MAXCUT}$, and
and 3SAT $\leq_p$ LONGPATH, all these problems are NP-hard as well. Finally, since all the aforementioned problems are in NP, they are all in fact NP-complete and have equivalent complexity. There are thousands of other natural problems that are NP-complete as well. Finding a polynomial-time algorithm for any one of them will imply a polynomial-time algorithm for all of them.

**Figure 14.8**: An instance of the independent set problem obtained by applying the reductions NANDSAT $\leq_p$ 3NAND $\leq_p$ 3SAT $\leq_p$ ISAT starting with the xor5 NAND-CIRC program.

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**Lecture Recap**

- Many of the problems for which we don’t know polynomial-time algorithms are NP-complete, which means that finding a polynomial-time algorithm for one of them would imply a polynomial-time algorithm for all of them.
- It is conjectured that NP $\neq$ P which means that we believe that polynomial-time algorithms for these problems are not merely unknown but are nonexistent.
- While an NP-hardness result means for example that a full-fledged “textbook” solution to a problem such as MAX-CUT that is as clean and general as the algorithm for MIN-CUT probably does not exist, it does not mean that we need to give up whenever we see a MAX-CUT instance. Later in this course we will discuss several strategies to deal with NP-hardness, including average-case complexity and approximation algorithms.

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**14.7 Exercises**

**Exercise 14.1 — Poor man’s Ladner’s Theorem.** Prove that if there is no $n^{O(\log^2 n)}$ time algorithm for 3SAT then there is some $F \in$ NP such that $F \notin$ P and $F$ is not NP complete.\(^9\)

\(^9\) Hint: Use the function $F$ that on input a formula $\varphi$ and a string of the form $1^t$, outputs 1 if and only if $\varphi$ is satisfiable and $t = |\varphi| \log |\varphi|$. \(^9\)
Exercise 14.2 Define WSAT to be the following function: the input is a CNF formula $\varphi$ where each clause is the OR of one to three variables (without negations), and a number $k \in \mathbb{N}$. For example, the following formula can be used for a valid input to WSAT: $\varphi = (x_5 \lor x_2 \lor x_1) \land (x_1 \lor x_3 \lor x_0) \land (x_2 \lor x_4 \lor x_0)$. The output $\text{WSAT}(\varphi, k) = 1$ if and only if there exists a satisfying assignment to $\varphi$ in which exactly $k$ of the variables get the value 1. For example, for the formula above $\text{WSAT}(\varphi, 2) = 1$ since the assignment $(1, 1, 0, 0, 0, 0)$ satisfies all the clauses. However, $\text{WSAT}(\varphi, 1) = 0$ since there is no single variable appearing in all clauses.

Prove that WSAT is NP-complete.

Exercise 14.3 In the employee recruiting problem we are given a list of potential employees, each of which has some subset of $m$ potential skills, and a number $k$. We need to assemble a team of $k$ employees such that for every skill there would be one member of the team with this skill.

For example, if Alice has the skills “C programming”, “NAND programming” and “Solving Differential Equations”, Bob has the skills “C programming” and “Solving Differential Equations”, and Charlie has the skills “NAND programming” and “Coffee Brewing”, then if we want a team of two people that covers all the four skills, we would hire Alice and Charlie.

Define the function $\text{EMP}$ s.t. on input the skills $L$ of all potential employees (in the form of a sequence $L$ of $n$ lists $L_1, \ldots, L_n$, each containing distinct numbers between 0 and $m$), and a number $k$, $\text{EMP}(L, k) = 1$ if and only if there is a subset $S$ of $k$ potential employees such that for every skill $j$ in $[m]$, there is an employee in $S$ that has the skill $j$.

Prove that EMP is NP complete.

Exercise 14.4 — Balanced max cut. Prove that the “balanced variant” of the maximum cut problem is NP-complete, where this is defined as $\text{BMC} : \{0, 1\}^* \rightarrow \{0, 1\}$ where for every graph $G = (V, E)$ and $k \in \mathbb{N}$, $\text{BMC}(G, k) = 1$ if and only if there exists a cut $S$ in $G$ cutting at least $k$ edges such that $|S| = |V|/2$.

Exercise 14.5 — Regular expression intersection. Let $\text{MANYREGS}$ be the following function: On input a list of regular expressions $\exp_0, \ldots, \exp_m$ (represented as strings in some standard way), output 1 if and only if
there is a single string \( x \in \{0, 1\}^* \) that matches all of them. Prove that \( \text{MANYREGS} \) is \( \text{NP} \)-hard.

14.8 BIBLIOGRAPHICAL NOTES

Eugene Lawler’s quote on the “mystical power of twoness” was taken from the wonderful book “The Nature of Computation” by Moore and Mertens. See also this memorial essay on Lawler by Lenstra.

14.9 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

14.10 ACKNOWLEDGEMENTS
15

What if \( P = \text{NP} \)?

We have mentioned that the question of whether \( P = \text{NP} \), which is equivalent to whether there is a polynomial-time algorithm for 3SAT, is the great open question of Computer Science. But why is it so important? In this chapter, we will try to figure out the implications of such an algorithm.

First, let us get one qualm out of the way. Sometimes people say, “What if \( P = \text{NP} \) but the best algorithm for 3SAT takes \( n^{100} \) time?” Well, \( n^{100} \) is much larger than, say, \( 2^{\sqrt{n}} \) for any input shorter than \( 10^{60} \) bits, which is way, way larger than the world’s total storage capacity (estimated at a “mere” \( 10^{21} \) bits or about 200 exabytes at the time of this writing). So another way to phrase this question is to say, “what if the complexity of 3SAT is exponential for all inputs that we will ever en-

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1 Paul Erdős (1913-1996) was one of the most prolific mathematicians of all times. Though he was an atheist, Erdős often referred to “The Book” in which God keeps the most elegant proof of each mathematical theorem.

2 The \( k \)-th Ramsey number, denoted as \( R(k, k) \), is the smallest number \( n \) such that for every graph \( G \) on \( n \) vertices, both \( G \) and its complement contain a \( k \)-sized independent set. If \( P = \text{NP} \) then we can compute \( R(k, k) \) in time polynomial in \( 2^k \), while otherwise it can potentially take closer to \( 2^{2^k} \) steps.
counter, but then grows much smaller than that?” To me this sounds like the computer science equivalent of asking, “what if the laws of physics change completely once they are out of the range of our telescopes?” Sure, this is a valid possibility, but wondering about it does not sound like the most productive use of our time.

So, as the saying goes, we’ll keep an open mind, but not so open that our brains fall out, and assume from now on that:

• There is a mathematical god,

and

• She does not “pussyfoot around” or take “half measures”. If God decided to make $3\text{SAT}$ easy, then $3\text{SAT}$ will have a $10^6 \cdot n$ (or at worst $10^6 n^2$) -time algorithm (i.e., $3\text{SAT}$ will be in $TIME(cn)$ or $TIME(cn^2)$ for a not-too-large constant $c$). If she decided to make $3\text{SAT}$ hard, then for every $n \in \mathbb{N}$, $3\text{SAT}$ on $n$ variables cannot be solved by a NAND-CIRC program of fewer than $2^{10^{-6n}}$ lines.\(^3\)

So far, most of our evidence points to the latter possibility of $3\text{SAT}$ being exponentially hard, but we have not ruled out the former possibility either. In this chapter we will explore some of its consequences.

15.1 SEARCH-TO-DECISION REDUCTION

A priori, having a fast algorithm for $3\text{SAT}$ might not seem so impressive. Sure, it will allow us to decide the satisfiability of not just $3\text{CNF}$ formulas but also of quadratic equations, as well as find out whether there is a long path in a graph, and solve many other decision problems. But this is not typically what we want to do. It’s not enough to know if a formula is satisfiable— we want to discover the actual satisfying assignment. Similarly, it’s not enough to find out if a graph has a long path— we want to actually find the path.

It turns out that if we can solve these decision problems, we can solve the corresponding search problems as well:

**Theorem 15.1 — Search vs Decision.** Suppose that $P = NP$. Then for every polynomial-time algorithm $V$ and $a, b \in \mathbb{N}$, there is a polynomial-time algorithm $\text{FIND}_V$ such that for every $x \in \{0, 1\}^n$, if there exists $y \in \{0, 1\}^{anb}$ satisfying $V(xy) = 1$, then $\text{FIND}_V(x)$ finds some string $y'$ satisfying this condition.

\(^3\) Using the relations we’ve seen between $SIZE(T(n))$ and $TIME(T(n))$ (i.e., Theorem 12.10), $3\text{SAT} \notin SIZE(T(n))$ then it is also in $TIME(T(n)^\epsilon)$ for some constant $\epsilon$ that can be shown to be at least $1/5$. 

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To understand what the statement of Theorem 15.1 means, let us look at the special case of the MAXCUT problem. It is not hard to see that there
is a polynomial-time algorithm \( VERIFYCUT \) such that \( VERIFYCUT(G, k, S) = 1 \) if and only if \( S \) is a subset of \( G \)'s vertices that cuts at least \( k \) edges. Theorem 15.1 implies that if \( P = NP \) then there is a polynomial-time algorithm \( FINDCUT \) that on input \( G, k \) outputs a set \( S \) such that \( VERIFYCUT(G, k, S) = 1 \) if such a set exists. This means that if \( P = NP \), by trying all values of \( k \) we can find in polynomial time a maximum cut in any given graph. We can use a similar argument to show that if \( P = NP \) then we can find a satisfying assignment for every satisfiable 3CNF formula, find the longest path in a graph, solve integer programming, and so and so forth.

**Proof Idea:**

The idea behind the proof of Theorem 15.1 is simple; let us demonstrate it for the special case of 3SAT. (In fact, this case is not so “special”—since 3SAT is NP-complete, we can reduce the task of solving the search problem for \( \text{MAXCUT} \) or any other problem in NP to the task of solving it for 3SAT.) Suppose that \( P = NP \) and we are given a satisfiable 3CNF formula \( \varphi \), and we now want to find a satisfying assignment \( y \) for \( \varphi \). Define \( 3SAT_0(\varphi) \) to output 1 if there is a satisfying assignment \( y \) for \( \varphi \) such that its first bit is 0, and similarly define \( 3SAT_1(\varphi) = 1 \) if there is a satisfying assignment \( y \) with \( y_0 = 1 \). The key observation is that both \( 3SAT_0 \) and \( 3SAT_1 \) are in NP, and so if \( P = NP \) then we can compute them in polynomial time as well. Thus we can use this to find the first bit of the satisfying assignment. We can continue in this way to recover all the bits.

\[
\star
\]

**Proof of Theorem 15.1.** If \( P = NP \) then for every polynomial-time algorithm \( V \) and \( a, b \in \mathbb{N} \), there is a polynomial-time algorithm \( STARTSWITH_V \) that on input \( x \in \{0, 1\}^a \) and \( z \in \{0, 1\}^b \), outputs 1 if and only if there exists some \( y \in \{0, 1\}^{an+b} \) such that the first \( \ell \) bits of \( y \) are equal to \( z \) and \( V(xy) = 1 \). Indeed, we leave it as an exercise to verify that the \( STARTSWITH_V \) function is in NP and hence can be solved in polynomial time if \( P = NP \).

Now for any such polynomial-time \( V \) and \( a, b \in \mathbb{N} \), we can implement \( FIND_V(x) \) as follows:
Algorithm 15.2 — FIND\_V: Search to decision reduction.

INPUT: \( x \in \{0, 1\}^* \) OUTPUT: \( x \in \{0, 1\}^{an^b} \) s.t. \( V(xz) = 1 \), if such \( x \) exists.

For \( \ell = 0, \ldots , an^b - 1 \) Let \( b_0 \leftarrow STARTSWITH(V(xz_0 \cdots z_{\ell-1}0)) \).

Let \( b_1 \leftarrow STARTSWITH(V(xz_0 \cdots z_{\ell-1}1)) \).

If \( b_0 = b_1 = 0 \)

Return “no \( z \) exists” # Can’t extend \( xz_0 \cdots z_{\ell-1}0 \) to accepting input of \( V \) Endif

If \( b = 1 \)

\( z_{\ell} \leftarrow 0 \) # Can extend \( xz_0 \cdots x_{\ell-1} \) with \( 0 \) to accepting input Else \( z_{\ell} \leftarrow 1 \) # Can extend \( xz_0 \cdots x_{\ell-1} \) with \( 1 \) to accepting input Endfor

Return \( z_0, \ldots , z_{an^b-1} \)

To analyze Algorithm 15.2, note that it makes \( 2an^b \) invocations to \( STARTSWITH \), and hence if the latter is polynomial-time, then so is Algorithm 15.2

Now suppose that \( x \) is such that there exists some \( y \) satisfying \( V(xy) = 1 \). We claim that at every step \( \ell = 0, \ldots , an^b - 1 \), we maintain the invariant that there exists \( y \in \{0, 1\}^{an^b} \) whose first \( \ell \) bits are \( z \) s.t. \( V(xy) = 1 \). Note that this claim implies the theorem, since in particular it means that for \( \ell = an^b - 1 \), \( z \) satisfies \( V(xz) = 1 \).

We prove the claim by induction. For \( \ell = 0 \), this holds vacuously. Now for every \( \ell > 0 \), if the call \( STARTSWITH(V(xz_0 \cdots z_{\ell-1}0)) \) returns 1, then we are guaranteed the invariant by definition of \( STARTSWITH \). Now under our inductive hypothesis, there is \( y_0, \ldots , y_{an^b-1} \) such that \( P(xz_0, \ldots , z_{\ell-1}0, y_0, \ldots , y_{an^b-1}) = 1 \). If the call to \( STARTSWITH(V(xz_0 \cdots z_{\ell-1}0)) \) returns 0 then it must be the case that \( y_0 = 1 \), and hence when we set \( z_{\ell} = 1 \) we maintain the invariant.

\[ \square \]

15.2 Optimization

Theorem 15.1 allows us to find solutions for \( \text{NP} \) problems if \( \text{P} = \text{NP} \), but it is not immediately clear that we can find the optimal solution. For example, suppose that \( \text{P} = \text{NP} \), and you are given a graph \( G \). Can you find the longest simple path in \( G \) in polynomial time?

This is actually an excellent question for you to attempt on your own. That is, assuming \( \text{P} = \text{NP} \), give a polynomial-time algorithm that on input a graph \( G \), outputs a maximally long simple path in the graph \( G \).

It turns out the answer is Yes. The idea is simple: if \( \text{P} = \text{NP} \) then we can find out in polynomial time if an \( n \)-vertex graph \( G \) contains a simple path of length \( n \), and moreover, by Theorem 15.1, if \( G \) does contain such a path, then we can find it. (Can you see why?) If \( G \) does not contain a simple path of length \( n \), then we will check if it contains
a simple path of length $n - 1$, and continue in this way to find the largest $k$ such that $G$ contains a simple path of length $k$.

The above reasoning was not specifically tailored to finding paths in graphs. In fact, it can be vastly generalized to proving the following result:

**Theorem 15.3 — Optimization from $P = NP$.** Suppose that $P = NP$. Then for every polynomial-time computable function $f : \{0, 1\}^* \rightarrow \{0, 1\}^*$ there is a polynomial-time algorithm $OPT$ such that on input $x \in \{0, 1\}^*$, $OPT(x, 1^m) = \max_{y \in \{0, 1\}^n} f(x, y)$ (where we identify the output of $f(x)$ with a natural number via the binary representation).

Moreover under the same assumption, there is a polynomial-time algorithm $FINDOPT$ such that for every $x \in \{0, 1\}^*$, $FINDOPT(x, 1^m)$ outputs $y^* \in \{0, 1\}^*$ such that $f(x, y^*) = OPT(x, y^*)$.

The statement of Theorem 15.3 is a bit cumbersome. To understand it, think how it would subsume the example above of a polynomial time algorithm for finding the maximum length path in a graph. In this case the function $f$ would be the map that on input a pair $x, y$ outputs 0 if the pair $(x, y)$ does not represent some graph and a simple path inside the graph respectively; otherwise $f(x, y)$ would equal the length of the path $y$ in the graph $x$. Since a path in an $n$ vertex graph can be represented by at most $n \log n$ bits, for every $x$ representing a graph of $n$ vertices, finding $\max_{y \in \{0, 1\}^n} f(x, y)$ corresponds to finding the length of the maximum simple path in the graph corresponding to $x$, and finding the string $y^*$ that achieves this maximum corresponds to actually finding the path.

**Proof Idea:**

The proof follows by generalizing our ideas from the longest path example above. Let $f$ be as in the theorem statement. If $P = NP$ then for every for every string $x \in \{0, 1\}^*$ and number $k$, we can test in $\text{poly}(|x|, m)$ time whether there exists $y$ such that $f(x, y) \geq k$, or in other words test whether $\max_{y \in \{0, 1\}^n} f(x, y) \geq k$. If $f(x, y)$ is an integer between 0 and $\text{poly}(|x| + |y|)$ (as is the case in the example of longest path) then we can just try out all possibilities for $k$ to find the maximum number $k$ for which $\max_y f(x, y) \geq k$. Otherwise, we can use binary search to hone down on the right value. Once we do so, we can use search-to-decision to actually find the string $y^*$ that achieves the maximum.
Proof of Theorem 15.3. For every \( f \) as in the theorem statement, we can define the Boolean function \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) as follows.

\[
F(x, 1^m, k) = \begin{cases} 
1 & \exists y \in \{0, 1\}^m \ f(x, y) \geq k \\
0 & \text{otherwise}
\end{cases}
\] (15.1)

Since \( f \) is computable in polynomial time, \( F \) is in \( \mathbf{NP} \), and so under our assumption that \( \mathbf{P} = \mathbf{NP} \), \( F \) itself can be computed in polynomial time. Now, for every \( x \) and \( m \), we can compute the largest \( k \) such that \( F(x, 1^m, k) = 1 \) by a binary search. Specifically, we will do this as follows:

1. We maintain two numbers \( a, b \) such that we are guaranteed that \( a \leq \max_{y \in \{0, 1\}^m} f(x, y) < b \).
2. Initially we set \( a = 0 \) and \( b = 2^T(n) \) where \( T(n) \) is the running time of \( f \). (A function with \( T(n) \) running time can’t output more than \( T(n) \) bits and so can’t output a number larger than \( 2^{2T(n)} \).)
3. At each point in time, we compute the midpoint \( c = \lfloor (a + b)/2 \rfloor \) and let \( y = F(1^n, c) \).
   a. If \( y = 1 \) then we set \( a = c \) and leave \( b \) as it is.
   b. If \( y = 0 \) then we set \( b = c \) and leave \( a \) as it is.
4. We then go back to step 3, until \( b \leq a + 1 \).

Since \( |b - a| \) shrinks by a factor of 2, within \( \log_2 2^{2T(n)} = T(n) \) steps, we will get to the point at which \( b \leq a + 1 \), and then we can simply output \( a \). Once we find the maximum value of \( k \) such that \( F(x, 1^m, k) = 1 \), we can use the search to decision reduction of Theorem 15.1 to obtain the actual value \( y^* \in \{0, 1\}^m \) such that \( f(x, y^*) = k \).

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Example 15.4 — Integer programming. One application for Theorem 15.3 is in solving optimization problems. For example, the task of linear programming is to find \( y \in \mathbb{R}^n \) that maximizes some linear objective \( \sum_{i=0}^{n-1} c_i y_i \) subject to the constraint that \( y \) satisfies linear inequalities of the form \( \sum_{i=0}^{n-1} a_i y_i \leq c \). As we discussed in Section 11.1.3, there is a known polynomial-time algorithm for linear programming. However, if we want to place additional constraints on \( y \), such as requiring the coordinates of \( y \) to be integer or 0/1 valued then the best-known algorithms run in exponential time in the worst case. However, if \( \mathbf{P} = \mathbf{NP} \) then Theorem 15.3 tells us
that we would be able to solve all problems of this form in polynomial time. For every string $x$ that describes a set of constraints and objective, we will define a function $f$ such that if $y$ satisfies the constraints of $x$ then $f(x, y)$ is the value of the objective, and otherwise we set $f(x, y) = -M$ where $M$ is some large number. We can then use Theorem 15.3 to compute the $y$ that maximizes $f(x, y)$ and that will give us the assignment for the variables that satisfies our constraints and maximizes the objective. (If the computation results in $y$ such that $f(x, y) = -M$ then we can double $M$ and try again; if the true maximum objective is achieved by some string $y^*$, then eventually $M$ will be large enough so that $-M$ would be smaller than the objective achieved by $y^*$, and hence when we run procedure of Theorem 15.3 we would get a value larger than $-M$.)

Remark 15.5 — Need for binary search. In many examples, such as the case of finding longest path, we don’t need to use the binary search step in Theorem 15.3, and can simply enumerate over all possible values for $k$ until we find the correct one. One example where we do need to use this binary search step is in the case of the problem of finding a maximum length path in a weighted graph. This is the problem where $G$ is a weighted graph, and every edge of $G$ is given a weight which is a number between 0 and $2^k$. Theorem 15.3 shows that we can find the maximum-weight simple path in $G$ (i.e., simple path maximizing the sum of the weights of its edges) in time polynomial in the number of vertices and in $k$.

Beyond just this example there is a vast field of mathematical optimization that studies problems of the same form as in Theorem 15.3. In the context of optimization, $x$ typically denotes a set of constraints over some variables (that can be Boolean, integer, or real valued), $y$ encodes an assignment to these variables, and $f(x, y)$ is the value of some objective function that we want to maximize. Given that we don’t know efficient algorithms for NP complete problems, researchers in optimization research study special cases of functions $f$ (such as linear programming and semidefinite programming) where it is possible to optimize the value efficiently. Optimization is widely used in a great many scientific areas including: machine learning, engineering, economics and operations research.
15.2.1 Example: Supervised learning

One classical optimization task is *supervised learning*. In supervised learning we are given a list of examples \(x_0, x_1, \ldots, x_{m-1}\) (where we can think of each \(x_i\) as a string in \(\{0, 1\}^n\) for some \(n\)) and the labels for them \(y_0, \ldots, y_{m-1}\) (which we will think of simply bits, i.e., \(y_i \in \{0, 1\}\)). For example, we can think of the \(x_i\)’s as images of either dogs or cats, for which \(y_i = 1\) in the former case and \(y_i = 0\) in the latter case. Our goal is to come up with a hypothesis or predictor \(h : \{0, 1\}^n \to \{0, 1\}\) such that if we are given a new example \(x\) that has an (unknown to us) label \(y\), then with high probability \(h\) will predict the label. That is, with high probability it will hold that \(h(x) = y\). The idea in supervised learning is to use the *Occam’s Razor principle*: the simplest hypothesis that explains the data is likely to be correct. There are several ways to model this, but one popular approach is to pick some fairly simple function \(H : \{0, 1\}^{k+n} \to \{0, 1\}\).

We think of the first \(k\) inputs as the *parameters* and the last \(n\) inputs as the example data. (For example, we can think of the first \(k\) inputs of \(H\) as specifying the weights and connections for some neural network that will then be applied on the latter \(n\) inputs.) We can then phrase the supervised learning problem as finding, given a set of labeled examples \(S = \{(x_0, y_0), \ldots, (x_{m-1}, y_{m-1})\}\), the set of parameters \(\theta_0, \ldots, \theta_{k-1} \in \{0, 1\}\) that minimizes the number of errors made by the predictor \(x \mapsto H(\theta, x)\).

In other words, we can define for every set \(S\) as above the function \(F_S : \{0, 1\}^k \to [m]\) such that \(F_S(\theta) = \sum_{(x, y) \in S} |H(\theta, x) - y|\). Now, finding the value \(\theta\) that minimizes \(F_S(\theta)\) is equivalent to solving the supervised learning problem with respect to \(H\). For every polynomial-time computable \(H : \{0, 1\}^{k+n} \to \{0, 1\}\), the task of minimizing \(F_S(\theta)\) can be “massaged” to fit the form of Theorem 15.3 and hence if \(P = NP\), then we can solve the supervised learning problem in great generality. In fact, this observation extends to essentially any learning model, and allows for finding the optimal predictors given the minimum number of examples. (This is in contrast to many current learning algorithms, which often rely on having access to an extremely large number of examples— far beyond the minimum needed, and in particular far beyond the number of examples humans use for the same tasks.)

15.2.2 Example: Breaking cryptosystems

We will discuss cryptography later in this course, but it turns out that if \(P = NP\) then almost every cryptosystem can be efficiently broken. One approach is to treat finding an encryption key as an instance of a supervised learning problem. If there is an encryption scheme that maps a “plaintext” message \(p\) and a key \(\theta\) to a “cipher-
text” \( c \), then given examples of ciphertext/plaintext pairs of the
form \((c_0, p_0), \ldots, (c_{m−1}, p_{m−1})\), our goal is to find the key \( \theta \) such that
\( E(\theta, p_i) = c_i \) where \( E \) is the encryption algorithm. While you might
think getting such “labeled examples” is unrealistic, it turns out (as
many amateur home-brew crypto designers learn the hard way) that
this is actually quite common in real-life scenarios, and that it is also
possible to relax the assumption to having more minimal prior information
about the plaintext (e.g., that it is English text). We defer a
more formal treatment to Chapter 20.

15.3 FINDING MATHEMATICAL PROOFS

In the context of Gödel’s Theorem, we discussed the notion of a proof
system (see Section 10.1). Generally speaking, a proof system can be
thought of as an algorithm \( V : \{0, 1\}^* \to \{0, 1\} \) (known as the verifier)
such that given a statement \( x \in \{0, 1\}^* \) and a candidate proof \( w \in \{0, 1\}^* \),
\( V(x, w) = 1 \) if and only if \( w \) encodes a valid proof for the statement \( x \).
Any type of proof system that is used in mathematics for geometry,
number theory, analysis, etc., is an instance of this form. In fact, standard
mathematical proof systems have an even simpler form where
the proof \( w \) encodes a sequence of lines \( w^0, \ldots, w^m \) (each of which is
itself a binary string) such that each line \( w^i \) is either an axiom or follows
from some prior lines through an application of some inference
rule. For example, Peano’s axioms encode a set of axioms and rules
for the natural numbers, and one can use them to formalize proofs
in number theory. Also, there are some even stronger axiomatic sys-
tems, the most popular one being Zermelo–Fraenkel with the Axiom
of Choice or ZFC for short. Thus, although mathematicians typically
write their papers in natural language, proofs of number theorists
can typically be translated to ZFC or similar systems, and so in par-
cular the existence of an \( n \)-page proof for a statement \( x \) implies that
there exists a string \( w \) of length \( \text{poly}(n) \) (in fact often \( O(n) \) or \( O(n^2) \))
that encodes the proof in such a system. Moreover, because verify-
ing a proof simply involves going over each line and checking that it
does indeed follow from the prior lines, it is fairly easy to do that in
\( O(|w|) \) or \( O(|w|^2) \) (where as usual \( |w| \) denotes the length of the proof
\( w \)). This means that for every reasonable proof system \( V \), the follow-
ing function \( \text{SHORTPROOF}_V : \{0, 1\}^* \to \{0, 1\} \) is in \( \text{NP} \), where
for every input of the form \( x1^m \), \( \text{SHORTPROOF}_V(x, 1^m) = 1 \) if and
only if there exists \( w \in \{0, 1\}^* \) with \( |w| \leq m \) s.t. \( V(xw) = 1 \). That
is, \( \text{SHORTPROOF}_V(x, 1^m) = 1 \) if there is a proof (in the system \( V \))
of length at most \( m \) bits that \( x \) is true. Thus, if \( P = \text{NP} \), then despite
Gödel’s Incompleteness Theorems, we can still automate mathematics
in the sense of finding proofs that are not too long for every statement
that has one. (Frankly speaking, if the shortest proof for some state-
The undecidability of the Entscheidungsproblem refers to the uncomputability of the function that maps a statement in first order logic to 1 if and only if that statement has a proof.

For many reasonable proof systems (including the one that Gödel referred to), SHORTPROOF$_V$ is in fact NP-complete, and so Gödel can be thought of as the first person to formulate the P vs NP question. Unfortunately, the letter was only discovered in 1988.

15.4 QUANTIFIER ELIMINATION (ADVANCED)

If P = NP then we can solve all NP search and optimization problems in polynomial time. But can we do more? It turns out that the answer is that Yes we can!

An NP decision problem can be thought of as the task of deciding, given some string $x \in \{0, 1\}^*$ the truth of a statement of the form

$$\exists y \in \{0, 1\}^{\geq |x|} \; V(xy) = 1$$

for some polynomial-time algorithm $V$ and polynomial $p : \mathbb{N} \to \mathbb{N}$. That is, we are trying to determine, given some string $x$, whether there exists a string $y$ such that $x$ and $y$ satisfy some polynomial-time checkable condition $V$. For example, in the independent set problem, the string $x$ represents a graph $G$ and a number $k$, the string $y$ represents some subset $S$ of $G$’s vertices, and the condition that we check is
what if $P = NP$?

Since NAND-CIRC programs are equivalent to Boolean circuits, the search problem corresponding to (15.5) known as the circuit minimization problem and is widely studied in Engineering. You can skip ahead to Section 15.4.1 to see a particularly compelling application of this.

For the ease of notation, we assume that all the strings we quantify over have the same length $m = p(n)$, but using simple padding one can show that this captures the general case of strings of different polynomial lengths.

Whether $|S| \geq k$ and there is no edge $\{u, v\}$ in $G$ such that both $u \in S$ and $v \in S$.

We can consider more general statements such as checking, given a string $x \in \{0, 1\}^*$, the truth of a statement of the form

$$\exists y \in \{0, 1\}^{p_0(x)} \forall z \in \{0, 1\}^{p_1(x)} V(xyz) = 1,$$

which in words corresponds to checking, given some string $x$, whether there exists a string $y$ such that for every string $z$, the triple $(x, y, z)$ satisfy some polynomial-time checkable condition. We can also consider more levels of quantifiers such as checking the truth of the statement

$$\exists y \in \{0, 1\}^{p_0(x)} \forall z \in \{0, 1\}^{p_1(x)} \exists w \in \{0, 1\}^{p_2(x)} V(xyzw) = 1$$

and so on and so forth.

For example, given an $n$-input NAND-CIRC program $P$, we might want to find the smallest NAND-CIRC program $P'$ that computes the same function as $P$. The question of whether there is such a $P'$ that can be described by a string of at most $s$ bits can be phrased as

$$\exists P' \in \{0, 1\}^{s} \forall x \in \{0, 1\}^{n} P(x) = P'(x)$$

which has the form (15.3). Another example of a statement involving $a$ levels of quantifiers would be to check, given a chess position $x$, whether there is a strategy that guarantees that White wins within $a$ steps. For example is $a = 3$ we would want to check if given the board position $x$, there exists a move $y$ for White such that for every move $z$ for Black there exists a move $w$ for White that ends in a a checkmate.

It turns out that if $P = NP$ then we can solve these kinds of problems as well:

**Theorem 15.6 — Polynomial hierarchy collapse.** If $P = NP$ then for every $a \in \mathbb{N}$, polynomial $p : \mathbb{N} \rightarrow \mathbb{N}$ and polynomial-time algorithm $V$, there is a polynomial-time algorithm $SOLVE_{V, a}$ that on input $x \in \{0, 1\}^n$ returns 1 if and only if

$$\exists \exists y_0 \in \{0, 1\}^m \forall y_1 \in \{0, 1\}^m \cdots \exists y_{a-1} \in \{0, 1\}^m V(xy_0y_1 \cdots y_{a-1}) = 1$$

where $m = p(n)$ and $\exists$ is either $\exists$ or $\forall$ depending on whether $a$ is odd or even, respectively.

**Proof Idea:**

To understand the idea behind the proof, consider the special case where we want to decide, given $x \in \{0, 1\}^n$, whether for every $y \in \{0, 1\}^n$ there exists $z \in \{0, 1\}^n$ such that $V(xyz) = 1$. Consider the function $F$ such that $F(xy) = 1$ if there exists $z \in \{0, 1\}^n$ such that $V(xyz) = 1$. Since $V$ runs in polynomial-time $F \in NP$ and hence if

$\exists y \in \{0, 1\}^{p_0(x)} \forall z \in \{0, 1\}^{p_1(x)} V(xyz) = 1$...
P = NP, then there is an algorithm $V'$ that on input $x, y$ outputs 1 if and only if there exists $z \in \{0, 1\}^n$ such that $V(xyz) = 1$. Now we can see that the original statement we consider is true if and only if there exists $z \in \{0, 1\}^n$ such that $V'(xy) = 1$. Now we can see that the original statement we consider is true if and only if for every $y \in \{0, 1\}^n$, $V'(xy) = 1$, which means it is false if and only if the following condition $(*)$ holds: there exists some $y \in \{0, 1\}^n$ such that $V'(xy) = 0$. But for every $x \in \{0, 1\}^n$, the question of whether the condition $(*)$ is itself in NP (as we assumed $V'$ can be computed in polynomial time) and hence under the assumption that $P = NP$ we can determine in polynomial time whether the condition $(*)$, and hence our original statement, is true.

⋆

Proof of Theorem 15.6. We prove the theorem by induction. We assume that there is a polynomial-time algorithm $SOLVE_{V,a-1}$ that can solve the problem (15.6) for $a-1$ and use that to solve the problem for $a$. For $a = 1$, $SOLVE_{V,a-1}(x) = 1$ if and only if $V(x) = 1$ which is a polynomial-time computation since $V$ runs in polynomial time. For every $x, y_0$, define the statement $\varphi_{x,y_0}$ to be the following:

$$\varphi_{x,y_0} = \forall y_1 \in \{0,1\}^m \exists y_2 \in \{0,1\}^m \cdots \exists y_{a-1} \in \{0,1\}^m V(xy_0y_1 \cdots y_{a-1}) = 1 \quad (15.7)$$

By the definition of $SOLVE_{V,a}$, for every $x \in \{0,1\}^n$, our goal is that $SOLVE_{V,a}(x) = 1$ if and only if there exists $y_0 \in \{0,1\}^m$ such that $\varphi_{x,y_0}$ is true.

The negation of $\varphi_{x,y_0}$ is the statement

$$\overline{\varphi}_{x,y_0} = \exists y_1 \in \{0,1\}^m \forall y_2 \in \{0,1\}^m \cdots \exists y_{a-1} \in \{0,1\}^m V(xy_0y_1 \cdots y_{a-1}) = 0 \quad (15.8)$$

where $\overline{\exists}$ is $\exists$ if $\exists$ was $\forall$ and $\overline{\forall}$ is $\forall$ otherwise. (Please stop and verify that you understand why this is true, this is a generalization of the fact that if $\Psi$ is some logical condition then the negation of $\exists y \forall z \Psi(y,z)$ is $\forall y \exists z \neg \Psi(y,z)$.)

The crucial observation is that $\overline{\varphi}_{x,y_0}$ is exactly a statement of the form we consider with $a - 1$ quantifiers instead of $a$, and hence by our inductive hypothesis there is some polynomial time algorithm $\overline{S}$ that on input $xy_0$ outputs 1 if and only if $\overline{\varphi}_{x,y_0}$ is true. If we let $S$ be the algorithm that on input $x, y_0$ outputs $1 - \overline{S}(xy_0)$ then we see that $S$ outputs 1 if and only if $\varphi_{x,y_0}$ is true. Hence we can rephrase the original statement (15.6) as follows:

$$\exists y_0 \in \{0,1\}^m S(xy_0) = 1 \quad (15.9)$$

but since $S$ is a polynomial-time algorithm, Eq. (15.9) is clearly a statement in NP and hence under our assumption that $P = NP$ there is
a polynomial time algorithm that on input $x \in \{0, 1\}^n$, will determine if (15.9) is true and so also if the original statement (15.6) is true.

The algorithm of Theorem 15.6 can also solve the search problem as well: find the value $y_0$ that certifies the truth of (15.6). We note that while this algorithm is in polynomial time, the exponent of this polynomial blows up quite fast. If the original NANDSAT algorithm required $\Omega(n^2)$ time, solving $a$ levels of quantifiers would require time $\Omega(n^{2^a})$.\(^8\)

15.4.1 Application: self improving algorithm for 3SAT
Suppose that we found a polynomial-time algorithm $A$ for 3SAT that is “good but not great”. For example, maybe our algorithm runs in time $cn^2$ for some not too small constant $c$. However, it’s possible that the best possible SAT algorithm is actually much more efficient than that. Perhaps, as we guessed before, there is a circuit $C^*$ of at most $10^6n$ gates that computes 3SAT on $n$ variables, and we simply haven’t discovered it yet. We can use Theorem 15.6 to “bootstrap” our original “good but not great” 3SAT algorithm to discover the optimal one. The idea is that we can phrase the question of whether there exists a size $s$ circuit that computes 3SAT for all length $n$ inputs as follows: there exists a size $\leq s$ circuit $C$ such that for every formula $\varphi$ described by a string of length at most $n$, if $C(\varphi) = 1$ then there exists an assignment $x$ to the variables of $\varphi$ that satisfies it. One can see that this is a statement of the form (15.4) and hence if $P = NP$ we can solve it in polynomial time as well. We can therefore imagine investing huge computational resources in running $A$ one time to discover the circuit $C^*$ and then using $C^*$ for all further computation.

15.5 APPROXIMATING COUNTING PROBLEMS AND POSTERIOR SAMPLING (ADVANCED, OPTIONAL)
Given a NAND-CIRC program $P$, if $P = NP$ then we can find an input $x$ (if one exists) such that $P(x) = 1$. But what if there is more than one $x$ like that? Clearly we can’t efficiently output all such $x$’s; there might be exponentially many. But we can get an arbitrarily good multiplicative approximation (i.e., a $1 \pm \epsilon$ factor for arbitrarily small $\epsilon > 0$) for the number of such $x$’s, as well as output a (nearly) uniform member of this set. We defer the details to later in this course, when we learn about randomized computation. However, we state (without proof) the following theorem for now:

\(^8\) We do not know whether such loss is inherent. As far as we can tell, it’s possible that the quantified boolean formula problem has a linear-time algorithm. We will, however, see later in this course that it satisfies a notion known as PSPACE-hardness that is even stronger than NP-hardness.
Theorem 15.7 — Approximate counting if $P = NP$. Let $V : \{0, 1\}^* \rightarrow \{0, 1\}$ be some polynomial-time algorithm, and suppose that $P = NP$. 
Then there exists an algorithm $\text{COUNT}_V$ that on input $x, 1^m, \epsilon$, runs in time polynomial in $|x|, m, 1/\epsilon$ and outputs a number $K \in \{0, \ldots, 2^m\}$ such that

$$(1 - \epsilon)K \leq \left| \{y \in \{0, 1\}^m : V(xy) = 1\} \right| \leq (1 + \epsilon)K \quad (15.10)$$

That is, $K$ gives an approximation up to a factor of $1 \pm \epsilon$ for the number of witnesses for $x$ with respect to the verifying algorithm $V$.

Once again, to understand this theorem it can be useful to see how it implies that if $P = \text{NP}$ then there is a polynomial-time algorithm that given a graph $G$ and a number $k$, can compute a number $K$ that is within a $1 \pm 0.01$ factor equal to the number of simple paths in $G$ of length $k$. (That is, $K$ is between 0.99 to 1.01 times the number of such paths.)

**Posterior sampling and probabilistic programming.** The algorithm for counting can also be extended to *sampling* from a given posterior distribution. That is, if $C : \{0, 1\}^n \to \{0, 1\}^m$ is a Boolean circuit and $y \in \{0, 1\}^m$, then if $P = \text{NP}$ we can sample from (a close approximation of) the distribution of uniform $x \in \{0, 1\}^n$ conditioned on $C(x) = y$. This task is known as *posterior sampling* and is crucial for Bayesian data analysis. These days it is known how to achieve posterior sampling only for circuits $C$ of very special form, and even in these cases more often than not we do have guarantees on the quality of the sampling algorithm. The field of making inferences by sampling from posterior distribution specified by circuits or programs is known as probabilistic programming.

**15.6 What does all of this imply?**

So, what will happen if we have a $10^6 n$ algorithm for 3SAT? We have mentioned that NP-hard problems arise in many contexts, and indeed scientists, engineers, programmers and others routinely encounter such problems in their daily work. A better 3SAT algorithm will probably make their lives easier, but that is the wrong place to look for the most foundational consequences. Indeed, while the invention of electronic computers did of course make it easier to do calculations that people were already doing with mechanical devices and pen and paper, the main applications computers are used for today were not even imagined before their invention.
An exponentially faster algorithm for all NP problems would be no less radical an improvement (and indeed, in some sense would be more) than the computer itself, and it is as hard for us to imagine what it would imply as it was for Babbage to envision today’s world. For starters, such an algorithm would completely change the way we program computers. Since we could automatically find the "best" (in any measure we chose) program that achieves a certain task, we would not need to define how to achieve a task, but only specify tests as to what would be a good solution, and could also ensure that a program satisfies an exponential number of tests without actually running them.

The possibility that $P = NP$ is often described as “automating creativity”. There is something to that analogy, as we often think of a creative solution as one that is hard to discover but that, once the “spark” hits, is easy to verify. But there is also an element of hubris to that statement, implying that the most impressive consequence of such an algorithmic breakthrough will be that computers would succeed in doing something that humans already do today. In fact, creativity already is to a large extent automated or minimized (e.g., just see how much popular media content is mass-produced), and as in most professions we should expect to see the need for humans diminish with time even if $P \neq NP$.

Nevertheless, artificial intelligence, like many other fields, will clearly be greatly impacted by an efficient 3SAT algorithm. For example, it is clearly much easier to find a better Chess-playing algorithm when, given any algorithm $P$, you can find the smallest algorithm $P'$ that plays Chess better than $P$. Moreover, as we mentioned above, much of machine learning (and statistical reasoning in general) is about finding “simple” concepts that explain the observed data, and if $NP = P$, we could search for such concepts automatically for any notion of “simplicity” we see fit. In fact, we could even “skip the middle man” and do an automatic search for the learning algorithm with smallest generalization error. Ultimately the field of Artificial Intelligence is about trying to “shortcut” billions of years of evolution to obtain artificial programs that match (or beat) the performance of natural ones, and a fast algorithm for $NP$ would provide the ultimate shortcut.9

More generally, a faster algorithm for NP problems would be immensely useful in any field where one is faced with computational or quantitative problems—which is basically all fields of science, math, and engineering. This will not only help with concrete problems such as designing a better bridge, or finding a better drug, but also with addressing basic mysteries such as trying to find scientific theories or “laws of nature”. In a fascinating talk, physicist Nima Arkani-Hamed...
discusses the effort of finding scientific theories in much the same language as one would describe solving an \textbf{NP} problem, for which the solution is easy to verify or seems “inevitable”, once found, but that requires searching through a huge landscape of possibilities to reach, and that often can get “stuck” at local optima:

“the laws of nature have this amazing feeling of inevitability... which is associated with local perfection.”

“The classical picture of the world is the top of a local mountain in the space of ideas. And you go up to the top and it looks amazing up there and absolutely incredible. And you learn that there is a taller mountain out there. Find it, Mount Quantum.... they’re not smoothly connected ... you’ve got to make a jump to go from classical to quantum ... This also tells you why we have such major challenges in trying to extend our understanding of physics. We don’t have these knobs, and little wheels, and twiddles that we can turn. We have to learn how to make these jumps. And it is a tall order. And that’s why things are difficult.”

Finding an efficient algorithm for \textbf{NP} amounts to always being able to search through an exponential space and find not just the “local” mountain, but the tallest peak.

But perhaps more than any computational speedups, a fast algorithm for \textbf{NP} problems would bring about a \textit{new type of understanding}. In many of the areas where \textbf{NP}-completeness arises, it is not as much a barrier for solving computational problems as it is a barrier for obtaining “closed-form formulas” or other types of more constructive descriptions of the behavior of natural, biological, social and other systems. A better algorithm for \textbf{NP}, even if it is “merely” $2^\sqrt{n}$-time, seems to require obtaining a new way to understand these types of systems, whether it is characterizing Nash equilibria, spin-glass configurations, entangled quantum states, or any of the other questions where \textbf{NP} is currently a barrier for analytical understanding. Such new insights would be very fruitful regardless of their computational utility.

\section*{15.7 CAN $P \neq \text{NP}$ BE NEITHER TRUE NOR FALSE?}

The Continuum Hypothesis is a conjecture made by Georg Cantor in 1878, positing the non-existence of a certain type of infinite cardinality.\textsuperscript{10} This was considered one of the most important open problems in set theory, and settling its truth or falseness was the first problem put forward by Hilbert in the 1900 address we mentioned before. However, using the theories developed by Gödel and Turing, in 1963 Paul Cohen proved that both the Continuum Hypothesis and its

\textsuperscript{10} One way to phrase it is that for every infinite subset $S$ of the real numbers $\mathbb{R}$, either there is a one-to-one and onto function $f : S \to \mathbb{R}$ or there is a one-to-one and onto function $f : S \to \mathbb{N}$. 


negation are consistent with the standard axioms of set theory (i.e., the Zermelo-Fraenkel axioms + the Axiom of choice, or “ZFC” for short).¹¹

Today, many (though not all) mathematicians interpret this result as saying that the Continuum Hypothesis is neither true nor false, but rather is an axiomatic choice that we are free to make one way or the other. Could the same hold for \( P \neq \text{NP} \)?

In short, the answer is No. For example, suppose that we are trying to decide between the “3SAT is easy” conjecture (there is an \( 10^6 n \) time algorithm for 3SAT) and the “3SAT is hard” conjecture (for every \( n \), any NAND-CIRC program that solves \( n \) variable 3SAT takes \( 2^{10^{-6} n} \) lines). Then, since for \( n = 10^8 \), \( 2^{10^{-6} n} > 10^6 n \), this boils down to the finite question of deciding whether or not there is a \( 10^{13} \)-line NAND-CIRC program deciding 3SAT on formulas with \( 10^8 \) variables. If there is such a program then there is a finite proof of its existence, namely the approximately 1TB file describing the program, and for which the verification is the (finite in principle though infeasible in practice) process of checking that it succeeds on all inputs.¹² If there isn’t such a program, then there is also a finite proof of that, though any such proof would take longer since we would need to enumerate over all programs as well. Ultimately, since it boils down to a finite statement about bits and numbers; either the statement or its negation must follow from the standard axioms of arithmetic in a finite number of arithmetic steps. Thus, we cannot justify our ignorance in distinguishing between the “3SAT easy” and “3SAT hard” cases by claiming that this might be an inherently ill-defined question. Similar reasoning (with different numbers) applies to other variants of the \( P \) vs \( \text{NP} \) question. We note that in the case that 3SAT is hard, it may well be that there is no short proof of this fact using the standard axioms, and this is a question that people have been studying in various restricted forms of proof systems.

### 15.8 IS \( P = \text{NP} \) “IN PRACTICE”?

The fact that a problem is \( \text{NP} \)-hard means that we believe there is no efficient algorithm that solve it in the worst case. It does not, however, mean that every single instance of the problem is hard. For example, if all the clauses in a 3SAT instance \( \varphi \) contain the same variable \( x_i \) (possibly in negated form), then by guessing a value to \( x_i \) we can reduce \( \varphi \) to a 2SAT instance which can then be efficiently solved. Generalizations of this simple idea are used in “SAT solvers”, which are algorithms that have solved certain specific interesting SAT formulas with thousands of variables, despite the fact that we believe SAT to be exponentially hard in the worst case. Similarly, a lot of problems arising in economics and machine learning are \( \text{NP} \)-hard.¹³ And yet

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¹¹ Formally, what he proved is that if ZFC is consistent, then so is ZFC when we assume either the continuum hypothesis or its negation.

¹² This inefficiency is not necessarily inherent. Later in this course we may discuss results in program-checking, interactive proofs, and average-case complexity, that can be used for efficient verification of proofs of related statements. In contrast, the inefficiency of verifying failure of all programs could well be inherent.

¹³ Actually, the computational difficulty of problems in economics such as finding optimal (or any) equilibria is quite subtle. Some variants of such problems are \( \text{NP} \)-hard, while others have a certain “intermediate” complexity.
WHAT IF P = NP?

So, $P = \text{NP}$ would give us all kinds of fantastical outcomes. But we strongly suspect that $P \neq \text{NP}$, and moreover that there is no much-better-than-brute-force algorithm for 3SAT. If indeed that is the case, is it all bad news?

One might think that impossibility results, telling you that you cannot do something, is the kind of cloud that does not have a silver lining. But in fact, as we already alluded to before, it does. A hard (in a sufficiently strong sense) problem in $\text{NP}$ can be used to create a code that cannot be broken, a task that for thousands of years has been the dream of not just spies but of many scientists and mathematicians over the generations. But the complexity viewpoint turned out to yield much more than simple codes, achieving tasks that people had previously not even dared to dream of. These include the notion of public key cryptography, allowing two people to communicate securely without ever having exchanged a secret key; electronic cash, allowing private and secure transaction without a central authority; and secure multiparty computation, enabling parties to compute a joint function on private inputs without revealing any extra information about it. Also,

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14 Talk more about coping with NP hardness. Main two approaches are heuristics such as SAT solvers that succeed on some instances, and proxy measures such as mathematical relaxations that instead of solving problem $X$ (e.g., an integer program) solve program $X'$ (e.g., a linear program) that is related to that. Maybe give compressed sensing as an example, and least square minimization as a proxy for maximum a posteriori probability.
as we will see, computational hardness can be used to replace the role of randomness in many settings.

Furthermore, while it is often convenient to pretend that computational problems are simply handed to us, and that our job as computer scientists is to find the most efficient algorithm for them, this is not how things work in most computing applications. Typically even formulating the problem to solve is a highly non-trivial task. When we discover that the problem we want to solve is \textbf{NP}-hard, this might be a useful sign that we used the wrong formulation for it.

Beyond all these, the quest to understand computational hardness — including the discoveries of lower bounds for restricted computational models, as well as new types of reductions (such as those arising from “probabilistically checkable proofs”) — has already had surprising 	extit{positive} applications to problems in algorithm design, as well as in coding for both communication and storage. This is not surprising since, as we mentioned before, from group theory to the theory of relativity, the pursuit of impossibility results has often been one of the most fruitful enterprises of mankind.

\section*{Lecture Recap}

- The question of whether $P = NP$ is one of the most important and fascinating questions of computer science and science at large, touching on all fields of the natural and social sciences, as well as mathematics and engineering.
- Our current evidence and understanding supports the “SAT hard” scenario that there is no much-better-than-brute-force algorithm for 3SAT or many other \textbf{NP}-hard problems.
- We are very far from proving this, however. Researchers have studied proving lower bounds on the number of gates to compute explicit functions in \textit{restricted forms} of circuits, and have made some advances in this effort, along the way generating mathematical tools that have found other uses. However, we have made essentially no headway in proving lower bounds for \textit{general} models of computation such as NAND and NAND-TM programs. Indeed, we currently do not even know how to rule out the possibility that for every $n \in \mathbb{N}$, \textit{SAT} restricted to $n$-length inputs has a NAND-CIRC program of $10n$ lines (even though there exist $n$-input functions that require $2^n/(10n)$ lines to compute).
- Understanding how to cope with this computational intractability, and even benefit from it, comprises much of the research in theoretical computer science.
15.10 EXERCISES

15.11 BIBLIOGRAPHICAL NOTES

15.12 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

- Polynomial hierarchy hardness for circuit minimization and related problems, see for example this paper.

15.13 ACKNOWLEDGEMENTS
Space bounded computation

PLAN: Example of space bounded algorithms, importance of preserving space. The classes L and PSPACE, space hierarchy theorem, PSPACE=NPSPACE, constant space = regular languages.

16.1 LECTURE SUMMARY

16.2 EXERCISES

16.3 BIBLIOGRAPHICAL NOTES

16.4 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

16.5 ACKNOWLEDGEMENTS
IV

RANDOMIZED COMPUTATION
Before we show how to use randomness in algorithms, let us do a quick review of some basic notions in probability theory. This is not meant to replace a course on probability theory, and if you have not seen this material before, I highly recommend you look at additional resources to get up to speed. Fortunately, we will not need many of the advanced notions of probability theory, but, as we will see, even the so-called “simple” setting of tossing $n$ coins can lead to very subtle and interesting issues.

### 17.1 RANDOM COINS

The nature of randomness and probability is a topic of great philosophical, scientific and mathematical depth. Is there actual random-
ness in the world, or does it proceed in a deterministic clockwork fashion from some initial conditions set at the beginning of time? Does probability refer to our uncertainty of beliefs, or to the frequency of occurrences in repeated experiments? How can we define probability over infinite sets?

These are all important questions that have been studied and debated by scientists, mathematicians, statisticians and philosophers. Fortunately, we will not need to deal directly with these questions here. We will be mostly interested in the setting of tossing \( n \) random, unbiased and independent coins. Below we define the basic probabilistic objects of events and random variables when restricted to this setting. These can be defined for much more general probabilistic experiments or sample spaces, and later on we will briefly discuss how this can be done. However, the \( n \)-coin case is sufficient for almost everything we’ll need in this course.

If instead of “heads” and “tails” we encode the sides of each coin by “zero” and “one”, we can encode the result of tossing \( n \) coins as a string in \( \{0,1\}^n \). Each particular outcome \( x \in \{0,1\}^n \) is obtained with probability \( 2^{-n} \). For example, if we toss three coins, then we obtain each of the 8 outcomes 000, 001, 010, 011, 100, 101, 110, 111 with probability \( 2^{-3} = 1/8 \) (see also Fig. 17.1). We can describe the experiment of tossing \( n \) coins as choosing a string \( x \) uniformly at random from \( \{0,1\}^n \), and hence we’ll use the shorthand \( x \sim \{0,1\}^n \) for \( x \) that is chosen according to this experiment.

An event is simply a subset \( A \) of \( \{0,1\}^n \). The probability of \( A \), denoted by \( \Pr_{x\sim\{0,1\}^n}[A] \) (or \( \Pr[A] \) for short, when the sample space is understood from the context), is the probability that an \( x \) chosen uniformly at random will be contained in \( A \). Note that this is the same as \( |A|/2^n \) (where \( |A| \) as usual denotes the number of elements in the set \( A \)). For example, the probability that \( x \) has an even number of ones is \( \Pr[A] \) where \( A = \{ x : \sum_{i=0}^{n-1} x_i = 0 \mod 2 \} \). In the case \( n = 3 \), \( A = \{000, 011, 101, 110\} \), and hence \( \Pr[A] = \frac{4}{8} = \frac{1}{2} \). It turns out this is true for every \( n \):

\[
\Pr_{x\sim\{0,1\}^n} \left[ \sum_{i=0}^{n-1} x_i \text{ is even} \right] = \frac{1}{2} \quad (17.1)
\]

**Lemma 17.1**

\[
\Pr_{x\sim\{0,1\}^n} \left[ \sum_{i=0}^{n-1} x_i \text{ is even} \right] = 1/2
\]

\[
\Pr_{x\sim\{0,1\}^n} \left[ \sum_{i=0}^{n-1} x_i \text{ is even} \right] = 1/2
\]

**Proof of Lemma 17.1.** Let \( A = \{ x \in \{0,1\}^n : \sum_{i=0}^{n-1} x_i = 0 \mod 2 \} \). Since every \( x \) is obtained with probability \( 2^{-n} \), to show this we need to...
show that $|A| = 2^n/2 = 2^{n-1}$. For every $x_0, \ldots, x_{n-2}$, if $\sum_{i=0}^{n-2} x_i$ is even then $(x_0, \ldots, x_{n-1}, 0) \in A$ and $(x_0, \ldots, x_{n-1}, 1) \notin A$. Similarly, if $\sum_{i=0}^{n-2} x_i$ is odd then $(x_0, \ldots, x_{n-1}, 1) \in A$ and $(x_0, \ldots, x_{n-1}, 0) \notin A$. Hence, for every one of the $2^{n-1}$ prefixes $(x_0, \ldots, x_{n-2})$, there is exactly a single continuation of $(x_0, \ldots, x_{n-2})$ that places it in $A$.

We can also use the intersection ($\cap$) and union ($\cup$) operators to talk about the probability of both event $A$ and event $B$ happening, or the probability of event $A$ or event $B$ happening. For example, the probability $p$ that $x$ has an even number of ones and $x_0 = 1$ is the same as $\Pr[A \cap B]$ where $A = \{x \in \{0, 1\}^n : \sum_{i=0}^{n-1} x_i = 0 \mod 2\}$ and $B = \{x \in \{0, 1\}^n : x_0 = 1\}$. This probability is equal to $1/4$ for $n > 1$. (It is a great exercise for you to pause here and verify that you understand why this is the case.)

Because intersection corresponds to considering the logical AND of the conditions that two events happen, while union corresponds to considering the logical OR, we will sometimes use the $\land$ and $\lor$ operators instead of $\cap$ and $\cup$, and so write this probability $p = \Pr[A \cap B]$ defined above also as

$$\Pr_{x \sim \{0,1\}^n} \left[ \sum_i x_i = 0 \mod 2 \land x_0 = 1 \right]. \quad (17.2)$$

If $A \subseteq \{0, 1\}^n$ is an event, then $\overline{A} = \{0, 1\}^n \setminus A$ corresponds to the event that $A$ does not happen. Since $|\overline{A}| = 2^n - |A|$, we get that

$$\Pr[\overline{A}] = \frac{|\overline{A}|}{2^n} = \frac{2^n - |A|}{2^n} = 1 - \frac{|A|}{2^n} = 1 - \Pr[A] \quad (17.3)$$

This makes sense: since $A$ happens if and only if $\overline{A}$ does not happen, the probability of $\overline{A}$ should be one minus the probability of $A$.

**Remark 17.2** — Remember the sample space. While the above definition might seem very simple and almost trivial, the human mind seems not to have evolved for probabilistic reasoning, and it is surprising how often people can get even the simplest settings of probability wrong. One way to make sure you don’t get confused when trying to calculate probability statements is to always ask yourself the following two questions: (1) Do I understand what is the sample space that this probability is taken over?, and (2) Do I understand what is the definition of the event that we are analyzing?

For example, suppose that I were to randomize seating in my course, and then it turned out that students sitting in row 7 performed better on the final: how surprising should we find this? If we started out with
the hypothesis that there is something special about
the number 7 and chose it ahead of time, then the
event that we are discussing is the event \( A \) that stu-
dents sitting in number 7 had better performance on
the final, and we might find it surprising. However, if
we first looked at the results and then chose the row
whose average performance is best, then the event
we are discussing is the event \( B \) that there exists some
row where the performance is higher than the over-
all average. \( B \) is a superset of \( A \), and its probability
(even if there is no correlation between sitting and
performance) can be quite significant.

17.1.1 Random variables

Events correspond to Yes/No questions, but often we want to analyze
finer questions. For example, if we make a bet at the roulette wheel,
we don’t want to just analyze whether we won or lost, but also how
much we’ve gained. A (real valued) random variable is simply a way
to associate a number with the result of a probabilistic experiment.
Formally, a random variable is simply a function \( X : \{0, 1\}^n \rightarrow \mathbb{R} \)
that maps every outcome \( x \in \{0, 1\}^n \) to a real number \( X(x) \).
For example, the function \( \text{sum} : \{0, 1\}^n \rightarrow \mathbb{R} \) that maps \( x \) to the sum of its
coordinates (i.e., to \( \sum_{i=0}^{n-1} x_i \)) is a random variable.

The expectation of a random variable \( X \), denoted by \( \mathbb{E}[X] \), is the
average value that that this number takes, taken over all draws from
the probabilistic experiment. In other words, the expectation of \( X \) is
defined as follows:

\[
\mathbb{E}[X] = \sum_{x \in \{0,1\}^n} 2^{-n} X(x). \tag{17.4}
\]

If \( X \) and \( Y \) are random variables, then we can define \( X + Y \) as
simply the random variable that maps a point \( x \in \{0, 1\}^n \) to \( X(x) +
Y(x) \). One basic and very useful property of the expectation is that it
is linear:

**Lemma 17.3 — Linearity of expectation.**

\[
\mathbb{E}[X + Y] = \mathbb{E}[X] + \mathbb{E}[Y] \tag{17.5}
\]

**Proof.**

\[
\mathbb{E}[X + Y] = \sum_{x \in \{0,1\}^n} 2^{-n} (X(x) + Y(x)) = \\
\sum_{x \in \{0,1\}^n} 2^{-n} X(x) + \sum_{x \in \{0,1\}^n} 2^{-n} Y(x) = \\
\mathbb{E}[X] + \mathbb{E}[Y] \tag{17.6}
\]

\[\Box\]
Similarly, $\mathbb{E}[kX] = k \mathbb{E}[X]$ for every $k \in \mathbb{R}$. For example, using the linearity of expectation, it is very easy to show that the expectation of the sum of the $x_i$’s for $x \sim \{0, 1\}^n$ is equal to $n/2$. Indeed, if we write $X = \sum_{i=0}^{n-1} x_i$ then $X = X_0 + \cdots + X_{n-1}$ where $X_i$ is the random variable $x_i$. Since for every $i$, $\Pr[X_i = 0] = 1/2$ and $\Pr[X_i = 1] = 1/2$, we get that $\mathbb{E}[X_i] = (1/2) \cdot 0 + (1/2) \cdot 1 = 1/2$ and hence $\mathbb{E}[X] = \sum_{i=0}^{n-1} \mathbb{E}[X_i] = n \cdot (1/2) = n/2$.

If $A$ is an event, then $1_A$ is the random variable such that $1_A(x)$ equals 1 if $x \in A$, and $1_A(x) = 0$ otherwise. Note that $\Pr[A] = \mathbb{E}[1_A]$ (can you see why?). Using this and the linearity of expectation, we can show one of the most useful bounds in probability theory:

**Lemma 17.4 — Union bound.** For every two events $A, B$, $\Pr[A \cup B] \leq \Pr[A] + \Pr[B]$

**Proof of Lemma 17.4.** For every $x$, the variable $1_{A \cup B}(x) \leq 1_A(x) + 1_B(x)$. Hence, $\Pr[A \cup B] = \mathbb{E}[1_{A \cup B}] \leq \mathbb{E}[1_A + 1_B] = \mathbb{E}[1_A] + \mathbb{E}[1_B] = \Pr[A] + \Pr[B]$.

The way we often use this in theoretical computer science is to argue that, for example, if there is a list of 100 bad events that can happen, and each one of them happens with probability at most $1/10000$, then with probability at least $1 - 100/10000 = 0.99$, no bad event happens.

**17.1.2 Distributions over strings**

While most of the time we think of random variables as having as output a real number, we sometimes consider random variables whose output is a string. That is, we can think of a map $Y : \{0, 1\}^n \to \{0, 1\}^*$ and consider the “random variable” $Y$ such

![Figure 17.2: The union bound tells us that the probability of $A$ or $B$ happening is at most the sum of the individual probabilities. We can see it by noting that for every two sets $|A \cup B| \leq |A| + |B|$ (with equality only if $A$ and $B$ have no intersection).](image-url)
that for every \( y \in \{0, 1\}^* \), the probability that \( Y \) outputs \( y \) is equal to \( \frac{1}{2^n} \cdot |\{ x \in \{0, 1\}^n \mid Y(x) = y \}| \). To avoid confusion, we will typically refer to such string-valued random variables as distributions over strings. So, a distribution \( Y \) over strings \( \{0, 1\}^* \) can be thought of as a finite collection of strings \( y_0, \ldots, y_{M-1} \in \{0, 1\}^* \) and probabilities \( p_0, \ldots, p_{M-1} \) (which are non-negative numbers summing up to one), so that \( \Pr[Y = y_i] = p_i \).

Two distributions \( Y \) and \( Y' \) are identical if they assign the same probability to every string. For example, consider the following two functions \( Y, Y' : \{0, 1\}^2 \rightarrow \{0, 1\}^2 \). For every \( x \in \{0, 1\}^2 \), we define \( Y(x) = x \) and \( Y'(x) = x_0(x_0 \oplus x_1) \) where \( \oplus \) is the XOR operation. Although these are two different functions, they induce the same distribution over \( \{0, 1\}^2 \) when invoked on a uniform input. The distribution \( Y(x) \) for \( x \sim \{0, 1\}^2 \) is of course the uniform distribution over \( \{0, 1\}^2 \). On the other hand \( Y' \) is simply the map \( 00 \mapsto 00, 01 \mapsto 01, 10 \mapsto 11, 11 \mapsto 10 \) which is a permutation over the map \( F : \{0, 1\}^2 \rightarrow \{0, 1\}^2 \) defined as \( F(x_0x_1) = x_0x_1 \) and the map \( G : \{0, 1\}^2 \rightarrow \{0, 1\}^2 \) defined as \( G(x_0x_1) = x_0(x_0 \oplus x_1) \).

### 17.1.3 More general sample spaces.

While in this chapter we assume that the underlying probabilistic experiment corresponds to tossing \( n \) independent coins, everything we say easily generalizes to sampling \( x \) from a more general finite or countable set \( S \) (and not-so-easily generalizes to uncountable sets \( S \) as well). A probability distribution over a finite set \( S \) is simply a function \( \mu : S \rightarrow [0, 1] \) such that \( \sum_{x \in S} \mu(x) = 1 \). We think of this as the experiment where we obtain every \( x \in S \) with probability \( \mu(s) \), and sometimes denote this as \( x \sim \mu \). An event \( A \) is a subset of \( S \), and the probability of \( A \), which we denote by \( \Pr_{\mu}[A] \), is \( \sum_{x \in A} \mu(x) \). A random variable is a function \( X : S \rightarrow \mathbb{R} \), where the probability that \( X = y \) is equal to \( \sum_{x \in S \text{ s.t. } X(x) = y} \mu(x) \).

### 17.2 CORRELATIONS AND INDEPENDENCE

One of the most delicate but important concepts in probability is the notion of independence (and the opposing notion of correlations). Subtle correlations are often behind surprises and errors in probability and statistical analysis, and several mistaken predictions have been blamed on miscalculating the correlations between, say, housing prices in Florida and Arizona, or voter preferences in Ohio and Michigan. See also Joe Blitzstein’s aptly named talk “Conditioning is the Soul of Statistics”.

Two events \( A \) and \( B \) are independent if the fact that \( A \) happens makes \( B \) neither more nor less likely to happen. For example, if we

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3 TODO: add exercise on simulating die tosses and choosing a random number in \([m]\) by coin tosses.

4 Another thorny issue is of course the difference between correlation and causation. Luckily, this is another point we don’t need to worry about in our clean setting of tossing \( n \) coins.
think of the experiment of tossing 3 random coins \( x \in \{0, 1\}^3 \), and we
let \( A \) be the event that \( x_0 = 1 \) and \( B \) the event that \( x_0 + x_1 + x_2 \geq 2 \),
then if \( A \) happens it is more likely that \( B \) happens, and hence these
events are not independent. On the other hand, if we let \( C \) be the event
that \( x_1 = 1 \), then because the second coin toss is not affected by the
result of the first one, the events \( A \) and \( C \) are independent.

The formal definition is that events \( A \) and \( B \) are independent if
\[
\Pr[A \cap B] = \Pr[A] \cdot \Pr[B].
\]
If \( \Pr[A \cap B] > \Pr[A] \cdot \Pr[B] \) then we say that
\( A \) and \( B \) are positively correlated, while if
\( \Pr[A \cap B] < \Pr[A] \cdot \Pr[B] \) then we say that \( A \) and \( B \) are negatively correlated (see Fig. 17.1).

If we consider the above examples on the experiment of choosing
\( x \in \{0, 1\}^3 \) then we can see that
\[
\Pr[x_0 = 1] = \frac{1}{2}
\]
\[
\Pr[x_0 + x_1 + x_2 \geq 2] = \Pr[\{011, 101, 110, 111\}] = \frac{4}{8} = \frac{1}{2}
\]
but
\[
\Pr[x_0 = 1 \land x_0 + x_1 + x_2 \geq 2] = \Pr[\{101, 110, 111\}] = \frac{3}{8} > \frac{1}{2} \cdot \frac{1}{2}
\]
and hence, as we already observed, the events \( \{x_0 = 1\} \) and \( \{x_0 +
x_1 + x_2 \geq 2\} \) are not independent and in fact are positively correlated.
On the other hand, \( \Pr[x_0 = 1 \land x_1 = 1] = \Pr[\{110, 111\}] = \frac{2}{8} = \frac{1}{2} \cdot \frac{1}{2} \)
and hence the events \( \{x_0 = 1\} \) and \( \{x_1 = 1\} \) are indeed independent.

**Remark 17.5 — Disjointness vs independence.** People
sometimes confuse the notion of disjointness and independence, but these are actually quite different. Two
events \( A \) and \( B \) are disjoint if \( A \cap B = \emptyset \), which means that if \( A \) happens then \( B \) definitely does not happen.
They are independent if \( \Pr[A \cap B] = \Pr[A] \Pr[B] \) which means that knowing that \( A \) happens gives us no information about whether \( B \) happened or not. If \( A \) and \( B \) have nonzero probability, then being disjoint implies that they are not independent, since in particular it
means that they are negatively correlated.

**Conditional probability:** If \( A \) and \( B \) are events, and \( A \) happens with
nonzero probability then we define the probability that \( B \) happens conditioned on \( A \) to be
\[
\Pr[B|A] = \frac{\Pr[A \cap B]}{\Pr[A]}. 
\]
This corresponds to calculating the probability that \( B \) happens if we already know
that \( A \) happened. Note that \( A \) and \( B \) are independent if and only if
\( \Pr[B|A] = \Pr[B] \).
More than two events: We can generalize this definition to more than two events. We say that events \( A_1, \ldots, A_k \) are mutually independent if knowing that any set of them occurred or didn’t occur does not change the probability that an event outside the set occurs. Formally, the condition is that for every subset \( I \subseteq \{k\}, \)

\[
\Pr[\bigwedge_{i \in I} A_i] = \prod_{i \in I} \Pr[A_i]. \tag{17.9}
\]

For example, if \( x \sim \{0, 1\}^3 \), then the events \( \{x_0 = 1\}, \{x_1 = 1\} \) and \( \{x_2 = 1\} \) are mutually independent. On the other hand, the events \( \{x_0 = 1\}, \{x_1 = 1\} \) and \( \{x_0 + x_1 = 0 \mod 2\} \) are not mutually independent, even though every pair of these events is independent (can you see why? see also Fig. 17.4).

17.2.1 Independent random variables

We say that two random variables \( X : \{0, 1\}^n \to \mathbb{R} \) and \( Y : \{0, 1\}^n \to \mathbb{R} \) are independent if for every \( u, v \in \mathbb{R} \), the events \( \{X = u\} \) and \( \{Y = v\} \) are independent.\(^5\) In other words, \( X \) and \( Y \) are independent if \( \Pr[X = u \land Y = v] = \Pr[X = u] \Pr[Y = v] \) for every \( u, v \in \mathbb{R} \). For example, if two random variables depend on the result of tossing different coins then they are independent:

**Lemma 17.6** Suppose that \( S = \{s_0, \ldots, s_{n-1}\} \) and \( T = \{t_0, \ldots, t_{m-1}\} \) are disjoint subsets of \( \{0, \ldots, n-1\} \) and let \( X, Y : \{0, 1\}^n \to \mathbb{R} \) be random variables such that \( X = F(x_{s_0}, \ldots, x_{s_{n-1}}) \) and \( Y = G(x_{t_0}, \ldots, x_{t_{m-1}}) \) for some functions \( F : \{0, 1\}^k \to \mathbb{R} \) and \( G : \{0, 1\}^m \to \mathbb{R} \). Then \( X \) and \( Y \) are independent.

**Proof of Lemma 17.6.** Let \( a, b \in \mathbb{R} \), and let \( A = \{ x \in \{0, 1\}^k : F(x) = a \} \) and \( B = \{ x \in \{0, 1\}^m : F(x) = b \} \). Since \( S \) and \( T \) are disjoint, we can reorder the indices so that \( S = \{0, \ldots, k-1\} \) and \( T = \{k, \ldots, k+m-1\} \) without affecting any of the probabilities. Hence we can write \( \Pr[X = a \land X = b] = \frac{|C|}{2^n} \) where \( C = \{x_0, \ldots, x_{n-1} : (x_0, \ldots, x_{k-1}) \in A \land (x_k, \ldots, x_{k+m-1}) \in B\} \). Another way to write this using string concatenation is that \( C = \{xyz : x \in A, y \in B, z \in \{0, 1\}^{n-k-m}\} \), and hence \( |C| = |A||B|2^{n-k-m} \), which means that

\[
\frac{|C|}{2^n} = \frac{|A||B|2^{n-k-m}}{2^n} = \Pr[X = a] \Pr[Y = b]. \tag{17.10}
\]

![Image](image.png)

**Figure 17.4:** Consider the sample space \( \{0, 1\}^n \) and the events \( A, B, C, D, E \) corresponding to \( A: x_0 = 1, B: x_1 = 1, C: x_0 + x_1 + x_2 \geq 2, D: x_0 + x_1 + x_2 = 0 \mod 2 \) and \( D: x_0 + x_1 = 0 \mod 2 \). We can see that \( A \) and \( B \) are independent, \( C \) is positively correlated with \( A \) and positively correlated with \( B \), the three events \( A, B, D \) are mutually independent, and while every pair out of \( A, B, E \) is independent, the three events \( A, B, E \) are not mutually independent since their intersection has probability \( \frac{2}{7} \) instead of \( \frac{1}{2} \).\(^5\) We use \( \{X = u\} \) as shorthand for \( \{x \mid X(x) = u\} \).
Note that if \( X \) and \( Y \) are independent random variables then (if we let \( S_X, S_Y \) denote all the numbers that have positive probability of being the output of \( X \) and \( Y \), respectively) it holds that:

\[
\mathbb{E}[XY] = \sum_{a \in S_X, b \in S_Y} \Pr[X = a \land Y = b] \cdot ab =^{(1)} \sum_{a \in S_X, b \in S_Y} \Pr[X = a] \Pr[Y = b] \cdot ab =^{(2)} \left( \sum_{a \in S_X} \Pr[X = a] \cdot a \right) \left( \sum_{b \in S_Y} \Pr[Y = b] \cdot b \right) =^{(3)} \mathbb{E}[X] \mathbb{E}[Y]
\]

where the first equality (\( =^{(1)} \)) follows from the independence of \( X \) and \( Y \), the second equality (\( =^{(2)} \)) follows by “opening the parentheses” of the righthand side, and the third inequality (\( =^{(3)} \)) follows from the definition of expectation. (This is not an “if and only if”; see Exercise 17.3.)

Another useful fact is that if \( X \) and \( Y \) are independent random variables, then so are \( F(X) \) and \( G(Y) \) for all functions \( F, G : \mathbb{R} \rightarrow \mathbb{R} \). This is intuitively true since learning \( F(X) \) can only provide us with less information than does learning \( X \) itself. Hence, if learning \( X \) does not teach us anything about \( Y \) (and so also about \( F(Y) \)) then neither will learning \( F(X) \). Indeed, to prove this we can write for every \( a, b \in \mathbb{R} \):

\[
\Pr[F(X) = a \land G(Y) = b] = \sum_{x \text{ s.t. } F(x) = a, y \text{ s.t. } G(y) = b} \Pr[X = x \land Y = y] = \\
\sum_{x \text{ s.t. } F(x) = a, y \text{ s.t. } G(y) = b} \Pr[X = x] \Pr[Y = y] = \\
\left( \sum_{x \text{ s.t. } F(x) = a} \Pr[X = x] \right) \cdot \left( \sum_{y \text{ s.t. } G(y) = b} \Pr[Y = y] \right) = \\
\Pr[F(X) = a] \Pr[G(Y) = b].
\]

(17.12)

### 17.2.2 Collections of independent random variables.

We can extend the notions of independence to more than two random variables: we say that the random variables \( X_0, \ldots, X_{n-1} \) are mutually independent if for every \( a_0, \ldots, a_{n-1} \in \mathbb{R} \),

\[
\Pr[X_0 = a_0 \land \cdots \land X_{n-1} = a_{n-1}] = \Pr[X_0 = a_0] \cdots \Pr[X_{n-1} = a_{n-1}].
\]

(17.13)

And similarly, we have that
Lemma 17.7 — Expectation of product of independent random variables. If $X_0, \ldots, X_{n-1}$ are mutually independent then

$$\mathbb{E} \left[ \prod_{i=0}^{n-1} X_i \right] = \prod_{i=0}^{n-1} \mathbb{E} [X_i]. \tag{17.14}$$

Lemma 17.8 — Functions preserve independence. If $X_0, \ldots, X_{n-1}$ are mutually independent, and $Y_0, \ldots, Y_{n-1}$ are defined as $Y_i = F_i(X_i)$ for some functions $F_0, \ldots, F_{n-1} : \mathbb{R} \to \mathbb{R}$, then $Y_0, \ldots, Y_{n-1}$ are mutually independent as well.

We leave proving Lemma 17.7 and Lemma 17.8 as Exercise 17.4 Exercise 17.5. It is good idea for you stop now and do these exercises to make sure you are comfortable with the notion of independence, as we will use it heavily later on in this course.

17.3 CONCENTRATION

The name “expectation” is somewhat misleading. For example, suppose that you and I place a bet on the outcome of 10 coin tosses, where if they all come out to be 1’s then I pay you 100,000 dollars and otherwise you pay me 10 dollars. If we let $X : \{0, 1\}^{10} \to \mathbb{R}$ be the random variable denoting your gain, then we see that

$$\mathbb{E} [X] = 2^{-10} \cdot 100000 - (1 - 2^{-10}) 10 \sim 90. \tag{17.15}$$

But we don’t really “expect” the result of this experiment to be for you to gain 90 dollars. Rather, 99.9% of the time you will pay me 10 dollars, and you will hit the jackpot 0.01% of the times.

However, if we repeat this experiment again and again (with fresh and hence independent coins), then in the long run we do expect your average earning to be 90 dollars, which is the reason why casinos can make money in a predictable way even though every individual bet is random. For example, if we toss $n$ coins, then as $n$ grows, the number of coins that come up ones will be more and more concentrated around $n/2$ according to the famous “bell curve” (see Fig. 17.5).

Much of probability theory is concerned with so called concentration or tail bounds, which are upper bounds on the probability that a random variable $X$ deviates too much from its expectation. The first and simplest one of them is Markov’s inequality:

**Theorem 17.9 — Markov’s inequality.** If $X$ is a non-negative random variable then $\Pr[X \geq \lambda \mathbb{E} [X]] \leq 1/\lambda$.

![Figure 17.5: The probabilities that we obtain a particular sum when we toss $n = 10, 20, 100, 1000$ coins converge quickly to the Gaussian/normal distribution.](image)
Markov’s Inequality is actually a very natural statement (see also Fig. 17.6). For example, if you know that the average (not the median!) household income in the US is 70,000 dollars, then in particular you can deduce that at most 25 percent of households make more than 280,000 dollars, since otherwise, even if the remaining 75 percent had zero income, the top 25 percent alone would cause the average income to be larger than 70,000. From this example you can already see that in many situations, Markov’s inequality will not be tight and the probability of deviating from expectation will be much smaller: see the Chebyshev and Chernoff inequalities below.

Proof of Theorem 17.9. Let \( \mu = \mathbb{E}[X] \) and define \( Y = 1_{X \geq k\mu} \). That is, \( Y(x) = 1 \) if \( X(x) \geq k\mu \) and \( Y(x) = 0 \) otherwise. Note that by definition, for every \( x \), \( Y(x) \leq X/(k\mu) \). We need to show \( \mathbb{E}[Y] \leq 1/k \).

But this follows since \( \mathbb{E}[Y] \leq \mathbb{E}[X/(k\mu)] = \mathbb{E}[X]/(k\mu) = \mu/(k\mu) = 1/k \).

**Going beyond Markov’s Inequality:** Markov’s inequality says that a (non-negative) random variable \( X \) can’t go too crazy and be, say, a million times its expectation, with significant probability. But ideally we would like to say that with high probability, \( X \) should be very close to its expectation, e.g., in the range \([0.99\mu, 1.01\mu]\) where \( \mu = \mathbb{E}[X] \). This is not generally true, but does turn out to hold when \( X \) is obtained by combining (e.g., adding) many independent random variables. This phenomenon, variants of which are known as “law of large numbers”, “central limit theorem”, “invariance principles” and “Chernoff bounds”, is one of the most fundamental in probability and statistics, and is one that we heavily use in computer science as well.

### 17.3.1 Chebyshev’s Inequality

A standard way to measure the deviation of a random variable from its expectation is by using its standard deviation. For a random variable \( X \), we define the variance of \( X \) as \( \text{Var}[X] = \mathbb{E}[(X - \mu)^2] \) where \( \mu = \mathbb{E}[X] \); i.e., the variance is the average squared distance of \( X \) from its expectation. The standard deviation of \( X \) is defined as \( \sigma[X] = \sqrt{\text{Var}[X]} \).

(This is well-defined since the variance, being an average of a square, is always a non-negative number.)

Using Chebyshev’s inequality, we can control the probability that a random variable is too many standard deviations away from its expectation.

---

Fig. 17.6: Markov’s Inequality tells us that a non-negative random variable \( X \) cannot be much larger than its expectation, with high probability. For example, if the expectation of \( X \) is \( \mu \), then the probability that \( X \) is at least \( 4\mu \) must be at most \( 1/4 \), as otherwise just the contribution from this part of the sample space will be too large.
Specifically, for a normal random variable \( X \) of expectation \( \mu \) and standard deviation \( \sigma \), the probability that \( |X - \mu| \geq k\sigma \) is at most \( 2e^{-k^2/2} \).

**Figure 17.7:** In the normal distribution or the bell curve, the probability of deviating \( k \) standard deviations from the expectation shrinks exponentially in \( k \), and specifically with probability at least \( 1 - 2e^{-k^2/2} \), a random variable \( X \) of expectation \( \mu \) and standard deviation \( \sigma \) satisfies \( \mu - k\sigma \leq X \leq \mu + k\sigma \). This figure gives more precise bounds for \( k = 1, 2, 3, 4, 5, 6 \).

(Exercise credit: Imran Baghirov)

**Theorem 17.10 — Chebyshev’s inequality.** Suppose that \( \mu = \mathbb{E}[X] \) and \( \sigma^2 = \text{Var}[X] \). Then for every \( k > 0 \), \( \Pr[|X - \mu| \geq k\sigma] \leq 1/k^2 \).

**Proof.** The proof follows from Markov’s inequality. We define the random variable \( Y = (X - \mu)^2 \). Then \( \mathbb{E}[Y] = \text{Var}[X] = \sigma^2 \), and hence by Markov the probability that \( Y \geq k^2\sigma^2 \) is at most \( 1/k^2 \). But clearly \( (X - \mu)^2 \geq k^2\sigma^2 \) if and only if \( |X - \mu| \geq k\sigma \).

One example of how to use Chebyshev’s inequality is the setting when \( X = X_1 + \cdots + X_n \) where the \( X_i \)’s are independent and identically distributed (i.i.d for short) variables with values in \([0, 1]\) where each has expectation \( 1/2 \). Since \( \mathbb{E}[X] = \sum_i \mathbb{E}[X_i] = n/2 \), we would like to say that \( X \) is very likely to be in, say, the interval \([0.499n, 0.501n]\). Using Markov’s inequality directly will not help us, since it will only tell us that \( X \) is very likely to be at most \( 100n \) (which we already knew, since it always lies between 0 and \( n \)). However, since \( X_1, \ldots, X_n \) are independent,

\[
\text{Var}[X_1 + \cdots + X_n] = \text{Var}[X_1] + \cdots + \text{Var}[X_n].
\]  

(17.16)

(We leave showing this to the reader as Exercise 17.6.)

For every random variable \( X_i \) in \([0, 1]\), \( \text{Var}[X_i] \leq 1 \) (if the variable is always in \([0, 1]\), it can’t be more than 1 away from its expectation), and hence (17.16) implies that \( \text{Var}[X] \leq n \) and hence \( \sigma[X] \leq \sqrt{n} \).

For large \( n, \sqrt{n} \ll 0.001n \), and in particular if \( \sqrt{n} \leq 0.001n/k \), we can use Chebyshev’s inequality to bound the probability that \( X \) is not in \([0.499n, 0.501n]\) by \( 1/k^2 \).

### 17.3.2 The Chernoff bound

Chebyshev’s inequality already shows a connection between independence and concentration, but in many cases we can hope for a quantitatively much stronger result. If, as in the example above, \( X = X_1 + \cdots + X_n \) where the \( X_i \)’s are bounded i.i.d random variables of mean \( 1/2 \), then as \( n \) grows, the distribution of \( X \) would be roughly the normal or Gaussian distribution— that is, distributed according to the bell curve (see Fig. 17.5 and Fig. 17.7). This distribution has the property of being very concentrated in the sense that the probability of deviating \( k \) standard deviations from the mean is not merely \( 1/k^2 \) as is guaranteed by Chebyshev, but rather is roughly \( e^{-k^2/2} \). That is, we have an exponential decay of the probability of deviation.

The following extremely useful theorem shows that such exponential decay occurs every time we have a sum of independent and bounded variables. This theorem is known under many names in different communities, though it is mostly called the Chernoff bound in the computer science literature:

\[\text{Specifically, for a normal random variable } X \text{ of expectation } \mu \text{ and standard deviation } \sigma, \text{ the probability that } |X - \mu| \geq k\sigma \text{ is at most } 2e^{-k^2/2}.\]
Theorem 17.11 — Chernoff/Hoeffding bound. If $X_1, \ldots, X_n$ are i.i.d random variables such that $X_i \in [0, 1]$ and $\mathbb{E}[X_i] = p$ for every $i$, then for every $\epsilon > 0$

$$\Pr\left[ \left| \sum_{i=0}^{n-1} X_i - pn \right| > \epsilon n \right] \leq 2 \cdot e^{-2\epsilon^2 n}. \quad (17.17)$$

We omit the proof, which appears in many texts, and uses Markov’s inequality on i.i.d random variables $Y_0, \ldots, Y_n$ that are of the form $Y_i = e^{\lambda X_i}$ for some carefully chosen parameter $\lambda$. See Exercise 17.9 for a proof of the simple (but highly useful and representative) case where each $X_i$ is $\{0, 1\}$ valued and $p = 1/2$. (See also Exercise 17.10 for a generalization.)

### 17.4 LECTURE SUMMARY

- A basic probabilistic experiment corresponds to tossing $n$ coins or choosing $x$ uniformly at random from $\{0, 1\}^n$.
- Random variables assign a real number to every result of a coin toss. The expectation of a random variable $X$ is its average value, and there are several concentration results showing that under certain conditions, random variables deviate significantly from their expectation only with small probability.

### 17.5 EXERCISES

**Exercise 17.1** Suppose that we toss three independent fair coins $a, b, c \in \{0, 1\}$. What is the probability that the XOR of $a, b, c$ is equal to 1? What is the probability that the AND of these three values is equal to 1? Are these two events independent?

**Exercise 17.2** Give an example of random variables $X, Y : \{0, 1\}^3 \to \mathbb{R}$ such that $\mathbb{E}[XY] \neq \mathbb{E}[X] \mathbb{E}[Y]$.

**Exercise 17.3** Give an example of random variables $X, Y : \{0, 1\}^3 \to \mathbb{R}$ such that $X$ and $Y$ are not independent but $\mathbb{E}[XY] = \mathbb{E}[X] \mathbb{E}[Y]$.

**Exercise 17.4 — Product of expectations.** Prove Lemma 17.7

**Exercise 17.5 — Transformations preserve independence.** Prove Lemma 17.8
Exercise 17.6 — Variance of independent random variables. Prove that if $X_0, \ldots, X_{n-1}$ are independent random variables then $\text{Var}[X_0 + \cdots + X_{n-1}] = \sum_{i=0}^{n-1} \text{Var}[X_i]$.

Exercise 17.7 — Entropy (challenge). Recall the definition of a distribution $\mu$ over some finite set $S$. Shannon defined the entropy of a distribution $\mu$, denoted by $H(\mu)$, to be $\sum_{x \in S} \mu(x) \log(1/\mu(x))$. The idea is that if $\mu$ is a distribution of entropy $k$, then encoding members of $\mu$ will require $k$ bits, in an amortized sense. In this exercise we justify this definition.

1. Prove that for every one to one function $F : S \to \{0, 1\}^*$, $\mathbb{E}_{x \sim \mu} |F(x)| \geq k$.
2. Prove that for every $\epsilon$, there is some $n$ and a one-to-one function $F : S^n \to \{0, 1\}^*$, such that $\mathbb{E}_{x \sim \mu^n} |F(x)| \leq n(k + \epsilon)$, where $x \sim \mu$ denotes the experiments of choosing $x_0, \ldots, x_{n-1}$ each independently from $S$ using the distribution $\mu$.

Exercise 17.8 — Entropy approximation to binomial. Let $H(p) = p \log(1/p) + (1-p) \log(1/(1-p))$. Prove that for every $p \in (0, 1)$ and $\epsilon > 0$, if $n$ is large enough then

$$2^{H(p) - \epsilon n} \binom{n}{k} \leq 2^{H(p) + \epsilon n}$$

where $\binom{n}{k}$ is the binomial coefficient $\frac{n!}{k!(n-k)!}$ which is equal to the number of $k$-size subsets of $\{0, \ldots, n-1\}$.

Exercise 17.9 — Chernoff using Stirling. 1. Prove that $\Pr_{x \sim \{0, 1\}^n}[\sum x_i = k] = \binom{n}{k}(2^{-n})$.
2. Use this and Exercise 17.8 to prove the Chernoff bound for the case that $X_0, \ldots, X_n$ are i.i.d. random variables over $\{0, 1\}$ each equaling 0 and 1 with probability 1/2.

Exercise 17.10 — Poor man’s Chernoff. Let $X_0, \ldots, X_n$ be i.i.d random variables with $\mathbb{E} X_i = p$ and $\Pr[0 \leq X_i \leq 1] = 1$. Define $Y_i = X_i - p$.

1. Prove that for every $j_1, \ldots, j_n \in \mathbb{N}$, if there exists one $i$ such that $j_i$ is odd then $\mathbb{E}[\prod_{i=0}^{n-1} Y_i^{j_i}] = 0$.
2. Prove that for every $k$, $\mathbb{E}[\sum_{i=0}^{n-1} Y_i^k] \leq (10kn)^{k/2}$.
3. Prove that for every $\epsilon > 0$, $\Pr[\sum Y_i \geq \epsilon n] \geq 2^{-\epsilon^2 n/(10000 \log 1/\epsilon)}$.

Exercise 17.11 — Simulating distributions using coins. Our model for probability involves tossing $n$ coins, but sometimes algorithms require sampling from other distributions, such as selecting a uniform number in $\{0, \ldots, M-1\}$ for some $M$. Fortunately, we can simulate this with an
exponentially small probability of error: prove that for every \( M \), if \( n > k\lceil \log M \rceil \), then there is a function \( F : \{0, 1\}^n \rightarrow \{0, \ldots, M - 1\} \cup \{\perp\} \) such that (1) The probability that \( F(x) = \perp \) is at most \( 2^{-k} \) and (2) the distribution of \( F(x) \) conditioned on \( F(x) \neq \perp \) is equal to the uniform distribution over \( \{0, \ldots, M - 1\} \).\(^{12}\)

**Exercise 17.12 — Sampling.** Suppose that a country has 300,000,000 citizens, 52 percent of which prefer the color “green” and 48 percent of which prefer the color “orange”. Suppose we sample \( n \) random citizens and ask them their favorite color (assume they will answer truthfully). What is the smallest value \( n \) among the following choices so that the probability that the majority of the sample answers “green” is at most 0.05?

a. 1,000    
b. 10,000     
c. 100,000    
d. 1,000,000

**Exercise 17.13** Would the answer to Exercise 17.12 change if the country had 300,000,000,000 citizens?

**Exercise 17.14 — Sampling (2).** Under the same assumptions as Exercise 17.12, what is the smallest value \( n \) among the following choices so that the probability that the majority of the sample answers “green” is at most \( 2^{-100} \)?

a. 1,000    
b. 10,000     
c. 100,000    
d. 1,000,000    
e. It is impossible to get such low probability since there are fewer than \( 2^{100} \) citizens.

\(^{12}\) **Hint:** Think of \( x \in \{0, 1\}^n \) as choosing \( k \) numbers \( y_1, \ldots, y_k \in \{0, \ldots, 2^\lceil \log M \rceil - 1\} \). Output the first such number that is in \( \{0, \ldots, M - 1\} \).

17.6 BIBLIOGRAPHICAL NOTES
18

Probabilistic computation

“In 1946 .. (I asked myself) what are the chances that a Canfield solitaire laid out with 52 cards will come out successfully? After spending a lot of time trying to estimate them by pure combinatorial calculations, I wondered whether a more practical method ... might not be to lay it our say one hundred times and simple observe and count”, Stanislaw Ulam, 1983

“The salient features of our method are that it is probabilistic ... and with a controllable miniscule probability of error.”, Michael Rabin, 1977

In early computer systems, much effort was taken to drive out randomness and noise. Hardware components were prone to non-deterministic behavior from a number of causes, whether it was vacuum tubes overheating or actual physical bugs causing short circuits (see Fig. 18.1). This motivated John von Neumann, one of the early computing pioneers, to write a paper on how to error correct computation, introducing the notion of redundancy.

So it is quite surprising that randomness turned out not just a hindrance but also a resource for computation, enabling us to achieve tasks much more efficiently than previously known. One of the first applications involved the very same John von Neumann. While he was sick in bed and playing cards, Stan Ulam came up with the observation that calculating statistics of a system could be done much faster by running several randomized simulations. He mentioned this idea to von Neumann, who became very excited about it; indeed, it turned out to be crucial for the neutron transport calculations that were needed for development of the Atom bomb and later on the hydrogen bomb. Because this project was highly classified, Ulam, von Neumann and their collaborators came up with the codeword “Monte Carlo” for this approach (based on the famous casinos where Ulam’s

Figure 18.1: A 1947 entry in the log book of the Harvard MARK II computer containing an actual bug that caused a hardware malfunction. By Courtesy of the Naval Surface Warfare Center.
uncle gambled). The name stuck, and randomized algorithms are known as Monte Carlo algorithms to this day.\footnote{Some texts also talk about “Las Vegas algorithms” that always return the right answer but whose running time is only polynomial on the average. Since this Monte Carlo vs Las Vegas terminology is confusing, we will not use these terms anymore, and simply talk about randomized algorithms.}

In this chapter, we will see some examples of randomized algorithms that use randomness to compute a quantity in a faster or simpler way than was known otherwise. We will describe the algorithms in an informal / “pseudo-code” way, rather than as NAND or NAND-TM programs. In Chapter 19 we will discuss how to augment the NAND and NAND-TM models to incorporate the ability to “toss coins”.

18.1 FINDING APPROXIMATELY GOOD MAXIMUM CUTS.

We start with the following example. Recall the maximum cut problem of finding, given a graph $G = (V, E)$, the cut that maximizes the number of edges. This problem is \textbf{NP}-hard, which means that we do not know of any efficient algorithm that can solve it, but randomization enables a simple algorithm that can cut at least half of the edges:

**Theorem 18.1 — Approximating max cut.** There is an efficient probabilistic algorithm that on input an $n$-vertex $m$-edge graph $G$, outputs a cut $(S, \overline{S})$ that cuts at least $m/2$ of the edges of $G$ in expectation.

**Proof Idea:**

We simply choose a random cut: we choose a subset $S$ of vertices by choosing every vertex $v$ to be a member of $S$ with probability $1/2$ independently. It’s not hard to see that each edge is cut with probability $1/2$ and so the expected number of cut edges is $m/2$.

$\star$

**Proof of Theorem 18.1.** The algorithm is extremely simple:

\begin{algorithm}
\caption{Random Cut:}
\begin{itemize}
  \item[Input:] Graph $G = (V, E)$ with $n$ vertices and $m$ edges. Denote $V = \{v_0, v_1, \ldots, v_{n-1}\}$.
  \item[Operation:]  
    \begin{enumerate}
      \item Pick $x$ uniformly at random in $\{0, 1\}^n$.
      \item Let $S \subseteq V$ be the set $\{v_i : x_i = 1 , i \in [n]\}$ that includes all vertices corresponding to coordinates of $x$ where $x_i = 1$.
      \item Output the cut $(S, \overline{S})$.
    \end{enumerate}
\end{itemize}
\end{algorithm}

We claim that the expected number of edges cut by the algorithm is $m/2$. Indeed, for every edge $e \in E$, let $X_e$ be the random variable such
that $X_e(x) = 1$ if the edge $e$ is cut by $x$, and $X_e(x) = 0$ otherwise. For every such edge $e = \{i, j\}$, $X_e(x) = 1$ if and only if $x_i \neq x_j$. Since the pair $(x_i, x_j)$ obtains each of the values 00, 01, 10, 11 with probability 1/4, the probability that $x_i \neq x_j$ is 1/2 and hence $E[X_e] = 1/2$. If we let $X$ be the random variable corresponding to the total number of edges cut by $S$, then $X = \sum_{e \in E} X_e$ and hence by linearity of expectation
\[ E[X] = \sum_{e \in E} E[X_e] = m(1/2) = m/2. \] (18.1)

18.1.1 Amplification

Theorem 18.1 gives us an algorithm that cuts $m/2$ edges in expectation. But, as we saw before, expectation does not immediately imply concentration, and so a priori, it may be the case that when we run the algorithm, most of the time we don’t get a cut matching the expectation. Luckily, we can amplify the probability of success by repeating the process several times and outputting the best cut we find. We start by arguing that the probability the algorithm above succeeds in cutting at least $m/2$ edges is not too tiny.

Lemma 18.2 The probability that a random cut in an $m$ edge graph cuts at least $m/2$ edges is at least $1/(2m)$.

Proof Idea:
To see the idea behind the proof, think of the case that $m = 1000$. In this case one can show that we will cut at least 500 edges with probability at least 0.001 (and so in particular larger than $1/(2m) = 1/2000$). Specifically, if we assume otherwise, then this means that with probability more than 0.999 the algorithm cuts 499 or fewer edges. But since we can never cut more than the total of 1000 edges, given this assumption, the highest value the expected number of edges cut is if we cut exactly 499 edges with probability 0.999 and cut 1000 edges with probability 0.001. Yet even in this case the expected number of edges will be $0.999 \cdot 499 + 0.001 \cdot 1000 < 500$, which contradicts the fact that we’ve calculated the expectation to be at least 500 in Theorem 18.1.

Proof of Lemma 18.2. Let $p$ be the probability that we cut at least $m/2$ edges and suppose, towards a contradiction, that $p < 1/(2m)$. Since the number of edges cut is an integer, and $m/2$ is a multiple of 0.5, by definition of $p$, with probability $1 - p$ we cut at most $m/2 - 0.5$ edges. Moreover, since we can never cut more than $m$ edges, under
our assumption that \( p < m/2 \), we can bound the expected number of edges cut by

\[
pm + (1 - p)(m/2 - 0.5) \leq pm + m/2 - 0.5 \quad (18.2)
\]

But if \( p < 1/(2m) \) then \( pm < 0.5 \) and so the righthand side is smaller than \( m/2 \), which contradicts the fact that (as proven in Theorem 18.1) the expected number of edges cut is at least \( m/2 \).

Success amplification. Lemma 18.2 shows that our algorithm succeeds at least some of the time, but we’d like to succeed almost all of the time. The approach to do that is to simply repeat our algorithm many times, with fresh randomness each time, and output the best cut we get in one of these repetitions. It turns out that with extremely high probability we will get a cut of size at least \( m/2 \). For example, if we repeat this experiment \( 2000m \) times, then (using the inequality \( (1 - 1/k)^k \leq 1/e \leq 1/2 \)) we can show that the probability that we will never cut at least \( m/2 \) edges is at most

\[
(1 - 1/(2m))^{2000m} \leq 2^{-1000} \quad (18.3)
\]

More generally, the same calculations can be used to show the following lemma:

**Lemma 18.3** There is a algorithm that on input a graph \( G = (V, E) \) and a number \( k \), runs in time polynomial in \( |V| \) and \( k \) and outputs a cut \((S, \overline{S})\) such that

\[
\Pr[\text{number of edges cut by } (S, \overline{S}) \geq |E|/2] \geq 1 - 2^{-k} \quad (18.4)
\]

**Proof of Lemma 18.3.** The algorithm will work as follows:

**Algorithm AMPLIFY RANDOM CUT:**

**Input:** Graph \( G = (V, E) \) with \( n \) vertices and \( m \) edges. Denote \( V = \{v_0, v_1, \ldots, v_{n-1}\} \). Number \( k > 0 \).

**Operation:**

1. Repeat the following \( 200km \) times:
   a. Pick \( x \) uniformly at random in \( \{0, 1\}^n \).
   b. Let \( S \subseteq V \) be the set \( \{v_i : x_i = 1, i \in [n]\} \) that includes all vertices corresponding to coordinates of \( x \) where \( x_i = 1 \).
   c. If \((S, \overline{S})\) cuts at least \( m/2 \) then halt and output \((S, \overline{S})\).
2. Output “failed”
We leave completing the analysis as an exercise to the reader (see Exercise 18.1).

### 18.1.2 Two-sided amplification

The analysis above relied on the fact that the maximum cut has one sided error. By this we mean that if we get a cut of size at least \(m/2\) then we know we have succeeded. This is common for randomized algorithms, but is not the only case. In particular, consider the task of computing some Boolean function \(F \colon \{0, 1\}^* \to \{0, 1\}\). A randomized algorithm \(A\) for computing \(F\), given input \(x\), might toss coins and succeed in outputting \(F(x)\) with probability, say, 0.9. We say that \(A\) has two sided errors if there is positive probability that \(A(x)\) outputs 1 when \(F(x) = 0\), and positive probability that \(A(x)\) outputs 0 when \(F(x) = 1\). In such a case, to simplify \(A\)'s success, we cannot simply repeat it \(k\) times and output 1 if a single one of those repetitions resulted in 1, nor can we output 0 if a single one of the repetitions resulted in 0. But we can output the majority value of these repetitions. By the Chernoff bound (Theorem 17.11), with probability exponentially close to 1 (i.e., \(1 - 2^{\Omega(k)}\)), the fraction of the repetitions where \(A\) will output \(F(x)\) will be at least, say 0.89, and in such cases we will of course output the correct answer.

The above translates into the following theorem

**Theorem 18.4 — Two-sided amplification.** If \(F \colon \{0, 1\}^* \to \{0, 1\}\) is a function such that there is a polynomial-time algorithm \(A\) satisfying

\[
\Pr[A(x) = F(x)] \geq 0.51 \tag{18.5}
\]

for every \(x \in \{0, 1\}^*\), then there is a polynomial time algorithm \(B\) satisfying

\[
\Pr[B(x) = F(x)] \geq 1 - 2^{-|x|} \tag{18.6}
\]

for every \(x \in \{0, 1\}^*\).

We omit the proof of Theorem 18.4, since we will prove a more general result later on in Theorem 19.4.

### 18.1.3 What does this mean?

We have shown a probabilistic algorithm that on any \(m\) edge graph \(G\), will output a cut of at least \(m/2\) edges with probability at least \(1 - 2^{-1000}\). Does it mean that we can consider this problem as “easy”? Should we be somewhat wary of using a probabilistic algorithm, since it can sometimes fail?

First of all, it is important to emphasize that this is still a worst case guarantee. That is, we are not assuming anything about the input.
introduction to theoretical computer science

This question does have some significance to practice, since hardware that generates high quality randomness at speed is nontrivial to construct.

\* The chance of winning the Massachusetts Mega Million lottery is one over \((75)^5 \cdot 15\) which is roughly \(2^{-35}\). So \(2^{-1000}\) corresponds to winning the lottery about 300 times in a row, at which point you might not care so much about your algorithm failing.

\* The chance for a U.S. resident to be struck by lightning is about \(1/700000\), which corresponds to about \(2^{-45}\) chance that you’ll be struck by lightning the very second that you’re reading this sentence (after which again you might not care so much about the algorithm’s performance).

\* Since the earth is about 5 billion years old, we can estimate the chance that an asteroid of the magnitude that caused the dinosaurs’ extinction will hit us this very second to be about \(2^{-60}\). It is quite likely that even a deterministic algorithm will fail if this happens.

So, in practical terms, a probabilistic algorithm is just as good as a deterministic one. But it is still a theoretically fascinating question whether randomized algorithms actually yield more power, or whether it is the case that for any computational problem that can be solved by probabilistic algorithm, there is a deterministic algorithm with nearly the same performance.\(^2\) For example, we will see in Exercise 18.2 that there is in fact a deterministic algorithm that can cut at least \(m/2\) edges in an \(m\)-edge graph. We will discuss this question in generality in Chapter 19. For now, let us see a couple of examples where randomization leads to algorithms that are better in some sense than the known deterministic algorithms.

18.1.4 Solving SAT through randomization

The 3SAT problem is \(\mathsf{NP}\) hard, and so it is unlikely that it has a polynomial (or even subexponential) time algorithm. But this does not mean that we can’t do at least somewhat better than the trivial \(2^n\) algorithm for \(n\)-variable 3SAT. The best known worst-case algorithms for 3SAT are randomized, and are related to the following simple algorithm, variants of which are also used in practice:

---

Algorithm WalkSAT:

\* Input: An \(n\) variable 3CNF formula \(\varphi\).

\* Parameters: \(T, S \in \mathbb{N}\)

\* Operation:

1. Repeat the following \(T\) steps:

---

\(^2\) This question does have some significance to practice, since hardware that generates high quality randomness at speed is nontrivial to construct.
At the time of this writing, the best known randomized algorithms for 3SAT run in time roughly $O(1.308^n)$, and the best known deterministic algorithms run in time $O(1.3303^n)$.  

The best algorithm we have is called WalkSAT.

The running time of this algorithm is $S \cdot T \cdot \text{poly}(n)$, and so the key question is how small we can make $S$ and $T$ so that the probability that WalkSAT outputs Unsatisfiable on a satisfiable formula $\varphi$ is small. It is known that we can do so with $ST = \tilde{O}((4/3)^n) = \tilde{O}(1.333...)$ (see Exercise 18.4 for a weaker result), but we’ll show below a simpler analysis yielding $ST = \tilde{O}(\sqrt{3} n) = \tilde{O}(1.74^n)$, which is still much better than the trivial $2^n$ bound.  

**Theorem 18.5 — WalkSAT simple analysis.** If we set $T = 100 \cdot \sqrt{3} n$ and $S = n/2$, then the probability we output Unsatisfiable for a satisfiable $\varphi$ is at most 1/2. 

**Proof.** Suppose that $\varphi$ is a satisfiable formula and let $x^*$ be a satisfying assignment for it. For every $x \in \{0,1\}^n$, denote by $\Delta(x,x^*)$ the number of coordinates that differ between $x$ and $x^*$. The heart of the proof is the following claim: 

**Claim I:** For every $x, x^*$ as above, in every local improvement step, the value $\Delta(x,x^*)$ is decreased by one with probability at least 1/3. 

**Proof of Claim I:** Since $x^*$ is a satisfying assignment, if $C$ is a clause that $x$ does not satisfy, then at least one of the variables involve in $C$ must get different values in $x$ and $x^*$. Thus when we change $x$ by one of the three literals in the clause, we have probability at least 1/3 of decreasing the distance. 

The second claim is that our starting point is not that bad: 

**Claim 2:** With probability at least 1/2 over a random $x \in \{0,1\}^n$, $\Delta(x,x^*) \leq n/2$. 

**Proof of Claim 2:** Consider the map $\text{FLIP} : \{0,1\}^n \rightarrow \{0,1\}^n$ that simply “flips” all the bits of its input from 0 to 1 and vice versa. That is, $\text{FLIP}(x_0, x_1, ..., x_{n-1}) = (1-x_0, 1-x_1, ..., 1-x_{n-1})$. Clearly $\text{FLIP}$ is one to one. Moreover, if $x$ is of distance $k$ to $x^*$, then $\text{FLIP}(x)$ is distance $n-k$ to $x^*$. Now let $B$ be the “bad event” in which $x$ is of distance $> n/2$ from $x^*$. Then the set $A = \text{FLIP}(B) = \{\text{FLIP}(x) : x \in \{0,1\}^n\}$ satisfies
\(|A| = |B|\) and that if \(x \in A\) then \(x\) is of distance \(< n/2\) from \(x^*\). Since \(A\) and \(B\) are disjoint events, \(\Pr[A] + \Pr[B] \leq 1\). Since they have the same cardinality, \(\Pr[A] + \Pr[B] \leq 1/2\). (See also Fig. 18.2).

Claims I and II imply that each of the \(T\) iterations of the outer loop succeeds with probability at least \(0.5 \cdot \sqrt{3}^{-n/2}\). Indeed, by Claim II, the original guess \(x\) will satisfy \(\Delta(x, x^*) \leq n/2\), and by Claim I, even conditioned on all the history so far, for each of the \(S = n/2\) steps we have probability \(\geq 1/3\) of being “lucky” and decreasing the distance at one. The chance we will be lucky in all \(n/2\) steps is hence at least \((1/3)^{n/2} = \sqrt{3}^{-n/2}\).

Since any single iteration of the outer loop succeeds with probability at least \(\frac{1}{2} \cdot \sqrt{3}^{-n}\), the probability that we never do so in \(T = 100\sqrt{3}^n\) repetitions is at most \((1 - \frac{1}{2\sqrt{3}})^{100\sqrt{3}^n} \leq (1/e)^{50}\).

18.1.5 Bipartite matching.
The matching problem is one of the canonical optimization problems, arising in all kinds of applications: matching residents with hospitals, kidney donors with patients, flights with crews, and many others.

One prototypical variant is bipartite perfect matching. In this problem, we are given a bipartite graph \(G = (L \cup R, E)\) which has \(2n\) vertices partitioned into \(n\)-sized sets \(L\) and \(R\), where all edges have one endpoint in \(L\) and the other in \(R\). The goal is to determine whether there is a perfect matching, a subset \(M \subseteq E\) of \(n\) disjoint edges. That is, \(M\) matches every vertex in \(L\) to a unique vertex in \(R\).

The bipartite matching problem turns out to have a polynomial-time algorithm, since we can reduce finding a matching in \(G\) to finding a maximum flow (or equivalently, minimum cut) in a related graph \(G'\) (see Fig. 18.3). However, we will see a different probabilistic algorithm to determine whether a graph contains such a matching.

Let us label \(G\)’s vertices as \(L = \{\ell_0, ..., \ell_{n-1}\}\) and \(R = \{r_0, ..., r_{n-1}\}\). A matching \(M\) corresponds to a permutation \(\pi \in S_n\) (i.e., one-to-one and onto function \(\pi : [n] \to [n]\) where for every \(i \in [n]\), we define \(\pi(i)\) to be the unique \(j\) such that \(M\) contains the edge \(\{\ell_i, r_j\}\). Define an \(n \times n\) matrix \(A = A(G)\) where \(A_{i,j} = 1\) if and only if the edge \(\{\ell_i, r_j\}\) is present and \(A_{i,j} = 0\) otherwise. The correspondence between matchings and permutations implies the following claim:

**Lemma 18.6 — Matching polynomial.** Define \(P = P(G)\) to be the polynomial mapping \(R^{n^2}\) to \(R\) where

\[
P(x_0,0, ..., x_{n-1,n-1}) = \sum_{\pi \in S_n} \left( \prod_{i=0}^{n-1} \text{sign}(\pi) A_{i,\pi(i)} \right) \prod_{i=0}^{n-1} x_{i,\pi(i)}
\]

Figure 18.2: For every \(x^* \in \{0, 1\}^n\), we can sort all strings in \(\{0, 1\}^n\) according to their distance from \(x^*\) (top to bottom in the above figure), where we let \(A = \{x \in \{0, 1\}^n \mid \text{dist}(x, x^*) \leq n/2\}\) be the “top half” of strings. If we define \(\text{FLIP} : \{0, 1\}^n \to \{0, 1\}^n\) to be the map that “flips” the bits of a given string \(x\) then it maps every \(x \in A\) to an output \(\text{FLIP}(x) \in A\) in a one-to-one way, and so it demonstrates that \(|\overline{A}| \leq |A|\) which implies that \(\Pr[A] \geq \Pr[\overline{A}]\) and hence \(\Pr[A] \geq 1/2\).

Figure 18.3: The bipartite matching problem in the graph \(G = (L \cup R, E)\) can be reduced to the minimum \(s, t\) cut problem in the graph \(G'\) obtained by adding vertices \(s, t\) to \(G\), connecting \(s\) with \(L\) and connecting \(t\) with \(R\).
Then $G$ has a perfect matching if and only if $P$ is not identically zero. That is, $G$ has a perfect matching if and only if there exists some assignment $x = (x_{i,j})_{i,j \in [n]} \in \mathbb{R}^{n^2}$ such that $P(x) \neq 0$.

**Proof.** If $G$ has a perfect matching $M^*$, then let $\pi^*$ be the permutation corresponding to $M$ and let $x^* \in \mathbb{Z}^{n^2}$ defined as follows: $x_{i,j} = 1$ if $j = \pi(i)$ and $x_{i,j} = 0$. Note that for every $\pi \neq \pi^*$, $\prod_{i=0}^{n-1} x_{i,\pi(i)} = 0$ but $\prod_{i=0}^{n-1} x_{i,\pi^*(i)} = 1$. Hence $P(x^*)$ will equal $\prod_{i=0}^{n-1} A_{i,\pi(i)}$. But since $M^*$ is a perfect matching in $G$, $\prod_{i=0}^{n-1} A_{i,\pi^*(i)} = 1$.

On the other hand, suppose that $P$ is not identically zero. By (18.7), this means that at least one of the terms $\prod_{i=0}^{n-1} A_{i,\pi(i)}$ is not equal to zero. But then this permutation $\pi$ must be a perfect matching in $G$.

As we’ve seen before, for every $x \in \mathbb{R}^{n^2}$, we can compute $P(x)$ by simply computing the determinant of the matrix $A(x)$, which is obtained by replacing $A_{i,j}$ with $A_{i,j}x_{i,j}$. This reduces testing perfect matching to the zero testing problem for polynomials: given some polynomial $P(\cdot)$, test whether $P$ is identically zero or not. The intuition behind our randomized algorithm for zero testing is the following:

*If a polynomial is not identically zero, then it can’t have “too many” roots.*

This intuition sort of makes sense. For one variable polynomials, we know that a nonzero linear function has at most one root, a quadratic function (e.g., a parabola) has at most two roots, and generally a degree $d$ equation has at most $d$ roots. While in more than one variable there can be an infinite number of roots (e.g., the polynomial $x_0 + y_0$ vanishes on the line $y = -x$) it is still the case that the set of roots is very “small” compared to the set of all inputs. For example, the root of a bivariate polynomial form a curve, the roots of a three-variable polynomial form a surface, and more generally the roots of an $n$-variable polynomial are a space of dimension $n - 1$.

This intuition leads to the following simple randomized algorithm:

*To decide if $P$ is identically zero, choose a “random” input $x$ and check if $P(x) \neq 0$.*

This makes sense: if there are only “few” roots, then we expect that with high probability the random input $x$ is not going to be one of those roots. However, to transform this into an actual algorithm, we need to make both the intuition and the notion of a “random” input precise. Choosing a random real number is quite problematic, especially when you have only a finite number of coins at your disposal, and so we start by reducing the task to a finite setting. We will use the following result:

4 The sign of a permutation $\pi : [n] \to [n]$, denoted by $\text{sign}(\pi)$, can be defined in several equivalent ways, one of which is that $\text{sign}(\pi) = (-1)^{\text{INV}(\pi)}$ where $\text{INV}(\pi) = \left| \{(x, y) \in [n] \mid x < y \land \pi(x) > \pi(y)\} \right|$ (i.e., $\text{INV}(\pi)$ is the number of pairs of elements that are inverted by $\pi$). The importance of the term $\text{sign}(\pi)$ is that it makes $P$ equal to the determinant of the matrix $(x_{i,j})$ and hence efficiently computable.
Theorem 18.7 — Schwartz–Zippel lemma. For every integer $q$, and polynomial $P : \mathbb{R}^n \to \mathbb{R}$ with integer coefficients. If $P$ has degree at most $d$ and is not identically zero, then it has at most $dq^{n-1}$ roots in the set $[q]^n = \{(x_0, \ldots, x_{n-1}) : x_i \in \{0, \ldots, q-1\}\}$.

We omit the (not too complicated) proof of Theorem 18.7. We remark that it holds not just over the real numbers but over any field as well. Since the matching polynomial $P$ of Lemma 18.6 has degree at most $n$, Theorem 18.7 leads directly to a simple algorithm for testing if it is nonzero:

**Algorithm Perfect-Matching:**

**Input:** Bipartite graph $G$ on $2n$ vertices
\{\ell_0, \ldots, \ell_{n-1}, r_0, \ldots, r_{n-1}\}.

**Operation:**

1. For every $i, j \in [n]$, choose $x_{i,j}$ independently at random from $[2n] = \{0, \ldots, 2n - 1\}$.
2. Compute the determinant of the matrix $A(x)$ whose $(i, j)^{th}$ entry equals $x_{i,j}$ if the edge $\{\ell_i, r_j\}$ is present and 0 otherwise.
3. Output no perfect matching if this determinant is zero, and output perfect matching otherwise.

This algorithm can be improved further (e.g., see Exercise 18.5). While it is not necessarily faster than the cut-based algorithms for perfect matching, it does have some advantages. In particular, it is more amenable for parallelization. (However, it also has the significant disadvantage that it does not produce a matching but only states that one exists.) The Schwartz–Zippel Lemma, and the associated zero testing algorithm for polynomials, is widely used across computer science, including in several settings where we have no known deterministic algorithm matching their performance.

**Lecture Recap**

- Using concentration results, we can amplify in polynomial time the success probability of a probabilistic algorithm from a mere $1/p(n)$ to $1 - 2^{-q(n)}$ for every polynomials $p$ and $q$.
- There are several randomized algorithms that are better in various senses (e.g., simpler, faster, or other advantages) than the best known deterministic algorithm for the same problem.
18.2 EXERCISES

Exercise 18.1 — Amplification for max cut. Prove Lemma 18.3

Exercise 18.2 — Deterministic max cut algorithm. 5

Exercise 18.3 — Simulating distributions using coins. Our model for probability involves tossing n coins, but sometimes algorithms require sampling from other distributions, such as selecting a uniform number in \{0, ..., M − 1\} for some M. Fortunately, we can simulate this with an exponentially small probability of error: prove that for every M, if \( n > k \lceil \log M \rceil \), then there is a function \( F : \{0, 1\}^n \rightarrow \{0, ..., M − 1\} \cup \{\perp\} \) such that (1) The probability that \( F(x) = \perp \) is at most \( 2^{-k} \) and (2) the distribution of \( F(x) \) conditioned on \( F(x) \neq \perp \) is equal to the uniform distribution over \{0, ..., M − 1\}. 6

Exercise 18.4 — Better walksat analysis. 1. Prove that for every \( \epsilon > 0 \), if \( n \) is large enough then for every \( x^* \in \{0, 1\}^n \) \( \Pr_{x \sim \{0, 1\}^n}[\Delta(x, x^*) \leq n/3] \leq 2^{-(1-H(1/3)-\epsilon)}n \) where \( H(p) = p \log(1/p) + (1-p) \log(1/(1-p)) \) is the same function as in Exercise 17.8.
2. Prove that \( 2^{1-H(1/4)+H(1/4)} \log 3 = (3/2) \).
3. Use the above to prove that for every \( \delta > 0 \) and large enough \( n \), if we set \( T = 1000 \cdot (3/2 + \delta)^n \) and \( S = n/4 \) in the WalkSAT algorithm then for every satisfiable 3CNF \( \varphi \), the probability that we output unsatisfiable is at most 1/2.

Exercise 18.5 — Faster bipartite matching (challenge). 7

18.3 BIBLIOGRAPHICAL NOTES


18.4 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

18.5 ACKNOWLEDGEMENTS
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Modeling randomized computation

“Any one who considers arithmetical methods of producing random digits is, of course, in a state of sin.” John von Neumann, 1951.

So far we have described randomized algorithms in an informal way, assuming that an operation such as “pick a string \( x \in \{0, 1\}^n \)” can be done efficiently. We have neglected to address two questions:

1. How do we actually efficiently obtain random strings in the physical world?

2. What is the mathematical model for randomized computations, and is it more powerful than deterministic computation?

The first question is of both practical and theoretical importance, but for now let’s just say that there are various physical sources of “random” or “unpredictable” data. A user’s mouse movements and typing pattern, (non solid state) hard drive and network latency, thermal noise, and radioactive decay have all been used as sources for randomness. For example, many Intel chips come with a random number generator built in. One can even build mechanical coin tossing machines (see Fig. 19.1).

In this chapter we focus on the second question: formally modeling probabilistic computation and studying its power. Modeling randomized computation is actually quite easy. We can add the following operations to our NAND, NAND-TM and NAND-RAM programming languages:

\[
\text{foo} = \text{RAND}()
\]

where \( \text{foo} \) is a variable. The result of applying this operation is that \( \text{foo} \) is assigned a random bit in \( \{0, 1\} \). (Every time the \text{RAND} operation is invoked it returns a fresh independent random bit.) We call

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Learning Objectives:
- Formal definition of probabilistic polynomial time: the class BPP.
- Proof that that every function in BPP can be computed by \( \text{poly}(n) \)-sized NAND-CIRC programs/circuits.
- Relations between BPP and NP.
- Pseudorandom generators

---

1 The output of processes such as above can be thought of as a binary string sampled from some distribution \( \mu \) that might have significant unpredictability (or entropy) but is not necessarily the uniform distribution over \( \{0, 1\}^n \). Indeed, as this paper shows, even (real-world) coin tosses do not have exactly the distribution of a uniformly random string. Therefore, to use the resulting measurements for randomized algorithms, one typically needs to apply a “distillation” or randomness extraction process to the raw measurements to transform them to the uniform distribution.

Figure 19.1: A mechanical coin tosser built for Percy Diaconis by Harvard technicians Steve Sansone and Rick Haggerty
the resulting languages RNAND, RNAND-TM, and RNAND-RAM respectively.

We can use this to define the notion of a function being computed by a randomized $T(n)$ time algorithm for every nice time bound $T : \mathbb{N} \to \mathbb{N}$, as well as the notion of a finite function being computed by a size $S$ randomized NAND-CIRC program (or, equivalently, a randomized circuit with $S$ gates that correspond to either the NAND or coin-tossing operations). However, for simplicity we will not define randomized computation in full generality, but simply focus on the class of functions that are computable by randomized algorithms running in polynomial time, which by historical convention is known as BPP:

Definition 19.1 — BPP. Let $F : \{0, 1\}^* \to \{0, 1\}$. We say that $F \in \text{BPP}$ if there exist constants $a, b \in \mathbb{N}$ and an RNAND-RAM program $P$ such that for every $x \in \{0, 1\}^*$, on input $x$, the program $P$ halts within at most $a|x|^b$ steps and

$$\Pr [P(x) = F(x)] \geq \frac{2}{3} \quad (19.1)$$

where this probability is taken over the result of the RAND operations of $P$.  

The same polynomial-overhead simulation of NAND-RAM programs by NAND-TM programs we saw in Theorem 12.5 extends to randomized programs as well. Hence the class BPP is the same regardless of whether it is defined via RNAND-TM or RNAND-RAM programs.

19.0.1 An alternative view: random coins as an “extra input”

While we presented randomized computation as adding an extra “coin tossing” operation to our programs, we can also model this as being given an additional extra input. That is, we can think of a randomized algorithm $A$ as a deterministic algorithm $A'$ that takes two inputs $x$ and $r$ where the second input $r$ is chosen at random from $\{0, 1\}^m$ for some $m \in \mathbb{N}$ (see Fig. 19.2). The equivalence to the Definition 19.1 is shown in the following theorem:

Theorem 19.2 — Alternative characterization of BPP. Let $F : \{0, 1\}^* \to \{0, 1\}$. Then $F \in \text{BPP}$ if and only if there exists $a, b \in \mathbb{N}$ and $G : \{0, 1\}^* \to \{0, 1\}$ such that $G$ is in P and for every $x \in \{0, 1\}^*$,

$$\Pr [G(xr) = F(x)] \geq \frac{2}{3} \quad (19.2)$$

Proof Idea:

BPP stands for “bounded probability polynomial time”, and is used for historical reasons.

![Figure 19.2: The two equivalent views of randomized algorithms.](image)
The idea behind the proof is that, as illustrated in Fig. 19.2, we can simply replace sampling a random coin with reading a bit from the extra ”random input” $r$ and vice versa. To prove this rigorously we need to work through some slightly cumbersome formal notation. This might be one of those proofs that is easier to work out on your own than to read.

\[\star\]

**Proof of Theorem 19.2.** We start by showing the “only if” direction. Let $F \in \text{BPP}$ and let $P$ be an RNAND-RAM program that computes $F$ as per Definition 19.1, and let $a, b \in \mathbb{N}$ be such that on every input of length $n$, the program $P$ halts within at most $a n^b$ steps. We will construct a polynomial-time NAND-RAM program $P'$ that computes a function $G$ satisfying the conditions of (19.2).

The program $P'$ is very simple:

**Program $P'$:** (Deterministic NAND-RAM program)

**Inputs:**
- $x \in \{0, 1\}^n$
- $r \in \{0, 1\}^{an^b}$

**Goal:** Output $y$ that has the same distribution as the output of the RNAND-RAM program $P$ on input $x$.

**Operation:**
1. Copy the string $r$ to an array $\text{Coins}$. That is $\text{Coins}[i] = r_i$ for all $i \in [an^b]$.
2. Let $\text{coincounter}$ be a variable and set it to 0.
3. Simulate an execution the RNAND-RAM program $P$, replacing any line of the form $\text{foo} = \text{RAND}()$ with the two lines: $\text{foo} = \text{Coins}[\text{coincounter}]$ and $\text{coincounter} = \text{coincounter} + 1$.

The program $P'$ is a deterministic polynomial time NAND-RAM program, and so computes some function $G \in \text{P}$. By its construction, the distribution of $P'(xr)$ for random $r \in \{0, 1\}^{2n^b}$ is identical to the distribution of $P(x)$ (where in the latter case the sample space is the results of the $\text{RAND}()$ calls), and hence in particular it will hold that $\Pr_{r \in \{0, 1\}^{an^b}}[P'(xr) = F(x)] \geq 2/3$.

For the other direction, given a function $G \in \text{P}$ satisfying the condition (19.2) and a NAND-RAM program $P'$ that computes $G$ in polynomial time, we can construct an RNAND-RAM program $P$ that computes $F$ in polynomial time. On input $x \in \{0, 1\}^n$, the program $P$
will simply use the RNAND() instruction $an^b$ times to fill an array $R[0], \ldots, R[an^b - 1]$ and then execute the original program $P'$ on input $xr$ where $r_i$ is the $i$-th element of the array $R$. Once again, it is clear that if $P'$ runs in polynomial time then so will $P$, and for every input $x$ and $r \in \{0, 1\}^an^b$, the output of $P$ on input $x$ and where the coin tosses outcome is $r$ is equal to $P'(xr)$.

Remark 19.3 — Definitions of BPP and NP. The characterization of BPP Theorem 19.2 is reminiscent of the characterization of NP in Definition 14.1, with the randomness in the case of BPP playing the role of the solution in the case of NP. However, there are important differences between the two:

- The definition of NP is "one sided": $F(x) = 1$ if there exists a solution $w$ such that $G(xw) = 1$ and $F(x) = 0$ if for every string $w$ of the appropriate length, $G(xw) = 0$. In contrast, the characterization of BPP is symmetric with respect to the cases $F(x) = 0$ and $F(x) = 1$.

- The relation between NP and BPP is not immediately clear. It is not known whether $\text{BPP} \subseteq \text{NP}$, $\text{NP} \subseteq \text{BPP}$, or these two classes are incomparable. It is however known (with a non-trivial proof) that if $P = \text{NP}$ then $\text{BPP} = P$ (see Theorem 19.11).

- Most importantly, the definition of NP is "ineffective," since it does not yield a way of actually finding whether there exists a solution among the exponentially many possibilities. By contrast, the definition of BPP gives us a way to compute the function in practice by simply choosing the second input at random.

Random tapes Theorem 19.2 motivates sometimes considering the randomness of an RNAND-TM (or RNAND-RAM) program as an extra input. As such, if $A$ is a randomized algorithm that on inputs of length $n$ makes at most $p(n)$ coin tosses, we will often use the notation $A(x; r)$ (where $x \in \{0, 1\}^n$ and $r \in \{0, 1\}^{p(n)}$) to refer to the result of executing $x$ when the coin tosses of $A$ correspond to the coordinates of $r$. This second, or "auxiliary," input is sometimes referred to as a "random tape." This terminology originates from the model of randomized Turing machines.

19.0.2 Amplification
The number $2/3$ might seem arbitrary, but as we’ve seen in Chapter 18 it can be amplified to our liking:
Theorem 19.4 — Amplification. Let \( P \) be an RNAND-RAM program, \( F \in \{0, 1\}^* \rightarrow \{0, 1\} \), and \( T : \mathbb{N} \rightarrow \mathbb{N} \) be a nice time bound such that for every \( x \in \{0, 1\}^* \), on input \( x \) the program \( P \) runs in at most \( T(|x|) \) steps and moreover \( \Pr[P(x) = F(x)] \geq \frac{1}{2} + \epsilon \) for some \( \epsilon > 0 \). Then for every \( k \), there is a program \( P' \) taking at most \( O(k \cdot T(n)/\epsilon^2) \) steps such that on input \( x \in \{0, 1\}^* \), \( \Pr[P'(x) = F(x)] > 1 - 2^{-k} \).

Proof Idea:

The proof is the same as we’ve seen before in the maximum cut and other examples. We use the Chernoff bound to argue that if we run the program \( O(k/\epsilon^2) \) times, each time using fresh and independent random coins, then the probability that the majority of the answers will not be correct will be less than \( 2^{-k} \). Amplification can be thought of as a “polling” of the choices for randomness for the algorithm (see Fig. 19.3).

Proof of Theorem 19.4. We can run \( P \) on input \( x \) for \( t = 10k/\epsilon^2 \) times, using fresh randomness in each execution, and compute the outputs \( y_0, \ldots, y_{t-1} \). We output the value \( y \) that appeared the largest number of times. Let \( X_i \) be the random variable that is equal to 1 if \( y_i = F(x) \) and equal to 0 otherwise. The random variables \( X_0, \ldots, X_{t-1} \) are i.i.d. and satisfy \( \mathbb{E}[X_i] = \Pr[X_i = 1] \geq 1/2 + \epsilon \), and hence by linearity of expectation \( \mathbb{E}[\sum_{i=0}^{t-1} X_i] \geq t(1/2 + \epsilon) \). For the plurality value to be incorrect, it must hold that \( \sum_{i=0}^{t-1} X_i \leq t/2 \), which means that \( \sum_{i=0}^{t-1} X_i \) is at least \( ct \) far from its expectation. Hence by the Chernoff bound (Theorem 17.11), the probability that the plurality value is not correct is at most \( 2e^{-\epsilon^2 t} \), which is smaller than \( 2^{-k} \) for our choice of \( t \).

There is nothing special about NAND-RAM in Theorem 19.4. The same proof can be used to amplify randomized NAND or NAND-TM programs as well.

19.1 BPP AND NP COMPLETENESS

Since “noisy processes” abound in nature, randomized algorithms can be realized physically, and so it is reasonable to propose \( \text{BPP} \) rather than \( \text{P} \) as our mathematical model for “feasible” or “tractable” computation. One might wonder if this makes all the previous chapters irrelevant, and in particular if the theory of \( \text{NP} \) completeness still applies to probabilistic algorithms. Fortunately, the answer is Yes:
**Theorem 19.5 — NP hardness and BPP.** Suppose that $F$ is NP-hard and $F \in \text{BPP}$. Then $\text{NP} \subseteq \text{BPP}$.

Before seeing the proof, note that Theorem 19.5 implies that if there was a randomized polynomial time algorithm for any NP-complete problem such as 3SAT, ISET etc., then there would be such an algorithm for every problem in NP. Thus, regardless of whether our model of computation is deterministic or randomized algorithms, NP complete problems retain their status as the “hardest problems in NP.”

**Proof Idea:**

The idea is to simply run the reduction as usual, and plug it into the randomized algorithm instead of a deterministic one. It would be an excellent exercise, and a way to reinforce the definitions of NP-hardness and randomized algorithms, for you to work out the proof for yourself. However for the sake of completeness, we include this proof below.

*Proof of Theorem 19.5. Suppose that $F$ is NP-hard and $F \in \text{BPP}$. We will now show that this implies that $\text{NP} \subseteq \text{BPP}$. Let $G \in \text{NP}$. By the definition of NP-hardness, it follows that $G \leq_p F$, or that in other words there exists a polynomial-time computable function $R : \{0, 1\}^* \rightarrow \{0, 1\}^*$ such that $G(x) = F(R(x))$ for every $x \in \{0, 1\}^*$. Now if $F$ is in BPP then there is a polynomial-time RNAND-TM program $P$ such that

$$\Pr[P(y) = F(y)] \geq 2/3$$

for every $y \in \{0, 1\}^*$ (where the probability is taken over the random coin tosses of $P$). Hence we can get a polynomial-time RNAND-TM program $P'$ to compute $G$ by setting $P'(x) = P(R(x))$. By (19.3) $$\Pr[P'(x) = F(R(x))] \geq 2/3$$ and since $F(R(x)) = G(x)$ this implies that $P'(x) = G(x)$, which proves that $G \in \text{BPP}$.*

Most of the results we’ve seen about NP hardness, including the search to decision reduction of Theorem 15.1, the decision to optimization reduction of Theorem 15.3, and the quantifier elimination result of Theorem 15.6, all carry over in the same way if we replace $P$ with BPP as our model of efficient computation. Thus if $\text{NP} \subseteq \text{BPP}$ then we get essentially all of the strange and wonderful consequences of $P = \text{NP}$. Unsurprisingly, we cannot rule out this possibility. In fact, unlike $P = \text{EXP}$, which is ruled out by the time hierarchy theorem, we don’t even know how to rule out the possibility that $\text{BPP} = \text{EXP}$! Thus a priori it’s possible (though seems highly unlikely) that randomness is a magical tool that allows us to speed up arbitrary exponential time
computation. Nevertheless, as we discuss below, it is believed that randomization’s power is much weaker and BPP lies in much more “pedestrian” territory.

19.2 THE POWER OF RANDOMIZATION

A major question is whether randomization can add power to computation. Mathematically, we can phrase this as the following question: does BPP = P? Given what we’ve seen so far about the relations of other complexity classes such as P and NP, or NP and EXP, one might guess that:

1. We do not know the answer to this question.
2. But we suspect that BPP is different than P.

One would be correct about the former, but wrong about the latter. As we will see, we do in fact have reasons to believe that BPP = P. This can be thought of as supporting the extended Church Turing hypothesis that deterministic polynomial-time NAND-TM program (or, equivalently, polynomial-time Turing machines) capture what can be feasibly computed in the physical world.

We now survey some of the relations that are known between BPP and other complexity classes we have encountered. (See also Fig. 19.4.)

19.2.1 Solving BPP in exponential time

It is not hard to see that if F is in BPP then it can be computed in exponential time.

Theorem 19.6 — Simulating randomized algorithms in exponential time.

\[ \text{BPP} \subseteq \text{EXP} \]

The proof of Theorem 19.6 readily follows by enumerating over all the (exponentially many) choices for the random coins. We omit the formal proof, as doing it by yourself is an excellent way to get comfortable with Definition 19.1.

19.2.2 Simulating randomized algorithms by circuits or straight-line programs.

We have seen in Theorem 12.10 that if F is in P, then there is a polynomial \( p : \mathbb{N} \to \mathbb{N} \) such that for every \( n \), the restriction \( F_{|n} \) of \( F \) to inputs \( \{0, 1\}^n \) is in \( \text{SIZE}(p(n)) \). (In other words, that \( P \subseteq \text{P/poly} \).) A priori it is
not at all clear that the same holds for a function in BPP, but this does turn out to be the case.

**Theorem 19.7 — Randomness does not help for non uniform computation.**

BPP ⊆ P/ poly. That is, for every $F \in$ BPP, there exist some $a, b \in \mathbb{N}$ such that for every $n > 0$, $F_n \in \text{SIZE}(an^b)$ where $F_n$ is the restriction of $F$ to inputs in $\{0, 1\}^n$.

**Proof Idea:**

The idea behind the proof is that we can first amplify by repetition the probability of success from $2/3$ to $1 - 0.1 \cdot 2^{-n}$. This will allow us to show that there exists a single fixed choice of “favorable coins” that would cause the algorithm to output the right answer on all of the possible $2^n$ inputs. We can then use the standard “unravelling the loop” technique to transform an RNAND-TM program to an RNAND-CIRC program, and “hardwire” the favorable choice of random coins to transform the RNAND-CIRC program into a plain old deterministic NAND-CIRC program.

*  

**Proof of Theorem 19.7.** Suppose that $F \in$ BPP. Let $P$ be a polynomial-time RNAND-TM program that computes $F$ as per Definition 19.1. Using Theorem 19.4, we can amplify the success probability of $P$ to obtain an RNAND-TM program $P'$ that is at most a factor of $O(n)$ slower (and hence still polynomial time) such that for every $x \in \{0, 1\}^n$

$$\Pr_{r \sim \{0, 1\}^m}[P'(x; r) = F(x)] \geq 1 - 0.1 \cdot 2^{-n}, \tag{19.4}$$

where $m$ is the number of coin tosses that $P'$ uses on inputs of length $n$. We use the notation $P'(x; r)$ to denote the execution of $P'$ on input $x$ and when the result of the coin tosses corresponds to the string $r$.

For every $x \in \{0, 1\}^n$, define the “bad” event $B_x$ to hold if $P'(x) \neq F(x)$, where the sample space for this event consists of the coins of $P'$. Then by (19.4), $\Pr[B_x] \leq 0.1 \cdot 2^{-n}$ for every $x \in \{0, 1\}^n$. Since there are $2^n$ many such $x$’s, by the union bound we see that the probability that the union of the events $\{B_x\}_{x \in \{0, 1\}^n}$ is at most 0.1. This means that if we choose $r \sim \{0, 1\}^m$, then with probability at least 0.9 it will be the case that for every $x \in \{0, 1\}^n$, $F(x) = P'(x; r)$. (Indeed, otherwise the event $B_x$ would hold for some $x$.) In particular, because of the mere fact that the the probability of $\bigcup_{x \in \{0, 1\}^n} B_x$ is smaller than 1, this means that there exists a particular $r^* \in \{0, 1\}^m$ such that

$$P'(x; r^*) = F(x) \tag{19.5}$$

![Figure 19.5: The possible guarantees for a randomized algorithm $A$ computing some function $F$. In the tables above, the columns correspond to different inputs and the rows to different choices of the random tape. A cell at position $r, x$ is colored green if $A(x; r) = F(x)$ (i.e., the algorithm outputs the correct answer) and red otherwise. The standard BPP guarantee corresponds to the middle figure, where for every input $x$, at least two thirds of the choices $r$ for a random tape will result in $A$ computing the correct value. That is, every column is colored green in at least two thirds of its coordinates. In the left figure we have an “average case” guarantee where the algorithm is only guaranteed to output the correct answer with probability two thirds over a random input (i.e., at most one third of the total entries of the table are colored red, but there could be an all red column). The right figure corresponds to the “offline BPP” case, with probability at least two thirds over the random choice $r$, $r$ will be good for every input. That is, at least two thirds of the rows are all green. Theorem 19.7 (BPP ⊆ P/ poly) is proven by amplifying the success of a BPP algorithm until we have the “offline BPP” guarantee, and then hardwiring the choice of the randomness $r$ to obtain a nonuniform deterministic algorithm.]

![Image of a table with green and red cells to illustrate the guarantees of a randomized algorithm.](image-url)
for every \( x \in \{0, 1\}^n \).

Now let us use the standard “unravelling the loop” the technique and transform \( P' \) into a NAND-CIRC program \( Q \) of polynomial in \( n \) size, such that \( Q(xr) = P'(x; r) \) for every \( x \in \{0, 1\}^n \) and \( r \in \{0, 1\}^m \). Then by “hardwiring” the values \( r_0, \ldots, r_{m-1} \) in place of the last \( m \) inputs of \( Q \), we obtain a new NAND-CIRC program \( Q_{r*} \) that satisfies by (19.5) that \( Q_{r*}(x) = F(x) \) for every \( x \in \{0, 1\}^n \). This demonstrates that \( F_{|n} \) has a polynomial sized NAND-CIRC program, hence completing the proof of Theorem 19.7.

---

**Remark 19.8 — Randomness and non uniformity.** The proof of Theorem 19.7 actually yields more than its statement. We can use the same “unrolling the loop” arguments we’ve used before to show that the restriction to \( \{0, 1\}^n \) of every function in \( \text{BPP} \) is also computable by a polynomial-size RNAND-CIRC program (i.e., NAND-CIRC program with the \( \text{RAND} \) operation). Like in the \( \text{P vs SIZE}(\text{poly}(n)) \) case, there are also functions outside \( \text{BPP} \) whose restrictions can be computed by polynomial-size RNAND-CIRC programs. Nevertheless the proof of Theorem 19.7 shows that even such functions can be computed by polynomial sized NAND-CIRC programs without using the \( r \) and operations. This can be phrased as saying that \( \text{BPSIZE}(T(n)) \subseteq \text{SIZE}(O(nT(n))) \) (where \( \text{BPSIZE} \) is defined in the natural way using RNAND programs). The stronger version of Theorem 19.7 we mentioned can be phrased as saying that \( \text{BPP}_{/\text{poly}} = \text{P}_{/\text{poly}} \).

### 19.3 Derandomization

The proof of Theorem 19.7 can be summarized as follows: we can replace a \( \text{poly}(n) \)-time algorithm that tosses coins as it runs with an algorithm that uses a single set of coin tosses \( r^* \in \{0, 1\}^{\text{poly}(n)} \) which will be good enough for all inputs of size \( n \). Another way to say it is that for the purposes of computing functions, we do not need “online” access to random coins and can generate a set of coins “offline” ahead of time, before we see the actual input.

But this does not really help us with answering the question of whether \( \text{BPP} \) equals \( \text{P} \), since we still need to find a way to generate these “offline” coins in the first place. To derandomize an RNAND-TM program we will need to come up with a single deterministic algorithm that will work for all input lengths. That is, unlike in the case of RNAND-CIRC programs, we cannot choose for every input length \( n \) some string \( r^* \in \{0, 1\}^{\text{poly}(n)} \) to use as our random coins.
Can we derandomize randomized algorithms, or does randomness add an inherent extra power for computation? This is a fundamentally interesting question but is also of practical significance. Ever since people started to use randomized algorithms during the Manhattan project, they have been trying to remove the need for randomness and replace it with numbers that are selected through some deterministic process. Throughout the years this approach has often been used successfully, though there have been a number of failures as well.

A common approach people used over the years was to replace the random coins of the algorithm by a “randomish looking” string that they generated through some arithmetic progress. For example, one can use the digits of \( \pi \) for the random tape. Using these type of methods corresponds to what von Neumann referred to as a “state of sin”. (Though this is a sin that he himself frequently committed, as generating true randomness in sufficient quantity was and still is often too expensive.) The reason that this is considered a “sin” is that such a procedure will not work in general. For example, it is easy to modify any probabilistic algorithm \( A \) such as the ones we have seen in Chapter 18, to an algorithm \( A' \) that is guaranteed to fail if the random tape happens to equal the digits of \( \pi \). This means that the procedure “replace the random tape by the digits of \( \pi \)” does not yield a general way to transform a probabilistic algorithm to a deterministic one that will solve the same problem. Of course, this procedure does not always fail, but we have no good way to determine when it fails and when it succeeds. This reasoning is not specific to \( \pi \) and holds for every deterministically produced string, whether it obtained by \( \pi \), \( e \), the Fibonacci series, or anything else.

An algorithm that checks if its random tape is equal to \( \pi \) and then fails seems to be quite silly, but this is but the “tip of the iceberg” for a very serious issue. Time and again people have learned the hard way that one needs to be very careful about producing random bits using deterministic means. As we will see when we discuss cryptography, many spectacular security failures and break-ins were the result of using “insufficiently random” coins.

### 19.3.1 Pseudorandom generators

So, we can’t use any single string to “derandomize” a probabilistic algorithm. It turns out however, that we can use a collection of strings to do so. Another way to think about it is that rather than trying to eliminate the need for randomness, we start by focusing on reducing the amount of randomness needed. (Though we will see that if we reduce the randomness sufficiently, we can eventually get rid of it altogether.)

We make the following definition:

---

4 One amusing anecdote is a recent case where scammers managed to predict the imperfect “pseudorandom generator” used by slot machines to cheat casinos. Unfortunately we don’t know the details of how they did it, since the case was sealed.
Definition 19.9 — Pseudorandom generator. A function $G : \{0, 1\}^\ell \rightarrow \{0, 1\}^m$ is a $(T, \epsilon)$-pseudorandom generator if for every NAND-CIRC program $P$ with $m$ inputs and one output of at most $T$ lines,

$$\left| \Pr_{s \sim \{0, 1\}^\ell} [P(G(s)) = 1] - \Pr_{r \sim \{0, 1\}^m} [P(r) = 1] \right| < \epsilon \tag{19.6}$$

This is a definition that’s worth reading more than once, and spending some time to digest it. Note that it takes several parameters:

- $T$ is the limit on the number of lines of the program $P$ that the generator needs to “fool”. The larger $T$ is, the stronger the generator.
- $\epsilon$ is how close is the output of the pseudorandom generator to the true uniform distribution over $\{0, 1\}^m$. The smaller $\epsilon$ is, the stronger the generator.
- $\ell$ is the input length and $m$ is the output length. If $\ell \geq m$ then it is trivial to come up with such a generator: on input $s \in \{0, 1\}^\ell$, we can output $s_0, \ldots, s_{m-1}$. In this case $\Pr_{s \sim \{0, 1\}^\ell} [P(G(s)) = 1]$ will simply equal $\Pr_{r \sim \{0, 1\}^m} [P(r) = 1]$, no matter how many lines $P$ has. So, the smaller $\ell$ is and the larger $m$ is, the stronger the generator, and to get anything non-trivial, we need $m > \ell$.

Furthermore note that although our eventual goal is to fool probabilistic randomized algorithms that take an unbounded number of inputs, Definition 19.9 refers to finite and deterministic NAND-CIRC programs.

We can think of a pseudorandom generator as a “randomness amplifier.” It takes an input $s$ of $\ell$ bits chosen at random and expands these $\ell$ bits into an output $r$ of $m > \ell$ pseudorandom bits. If $\epsilon$ is small enough then the pseudorandom bits will “look random” to any NAND-CIRC program that is not too big. Still, there are two questions we haven’t answered:

- What reason do we have to believe that pseudorandom generators with non-trivial parameters exist?

- Even if they do exist, why would such generators be useful to derandomize randomized algorithms? After all, Definition 19.9 does not involve RNAND-TM or RNAND-RAM programs, but rather deterministic NAND-CIRC programs with no randomness and no loops.
We will now (partially) answer both questions. For the first question, let us come clean and confess we do not know how to prove that interesting pseudorandom generators exist. By *interesting* we mean pseudorandom generators that satisfy that $\epsilon$ is some small constant (say $\epsilon < 1/3$), $m > \ell$, and the function $G$ itself can be computed in $\text{poly}(m)$ time. Nevertheless, Lemma 19.12 (whose statement and proof is deferred to the end of this chapter) shows that if we only drop the last condition (polynomial-time computability), then there do in fact exist pseudorandom generators where $m$ is exponentially larger than $\ell$. 

At this point you might want to skip ahead and look at the statement of Lemma 19.12. However, since its proof is somewhat subtle, I recommend you defer reading it until you’ve finished reading the rest of this chapter.

19.3.2 From existence to constructivity

The fact that there exists a pseudorandom generator does not mean that there is one that can be efficiently computed. However, it turns out that we can turn complexity “on its head” and use the assumed non existence of fast algorithms for problems such as 3SAT to obtain pseudorandom generators that can then be used to transform randomized algorithms into deterministic ones. This is known as the *Hardness vs Randomness* paradigm. A number of results along those lines, most of which are outside the scope of this course, have led researchers to believe the following conjecture:

**Optimal PRG conjecture:** There is a polynomial-time computable function $\text{PRG} : \{0, 1\}^* \rightarrow \{0, 1\}$ that yields an exponentially secure pseudorandom generator.

Specifically, there exists a constant $\delta > 0$ such that for every $\ell$ and $m < 2^\delta \ell$, if we define $G : \{0, 1\}^\ell \rightarrow \{0, 1\}^m$ as $G(s)_i = \text{PRG}(s, i)$ for every $s \in \{0, 1\}^\ell$ and $i \in [m]$, then $G$ is a $(2^\delta, 2^{-\delta})$ pseudorandom generator.

The “optimal PRG conjecture” is worth while reading more than once. What it posits is that we can obtain $(T, \epsilon)$ pseudorandom generator $G$ such that every output bit of $G$ can be computed in time polynomial in the length $\ell$ of the input, where $T$ is exponentially large in $\ell$ and $\epsilon$ is exponentially small in $\ell$. (Note that
we could not hope for the entire output to be computable in $\ell$, as just writing the output down will take too long.)

To understand why we call such a pseudorandom generator “optimal,” it is a great exercise to convince yourself that, for example, there does not exist a $(2^{1.1\ell}, 2^{-1.1\ell})$ pseudorandom generator (in fact, the number $\delta$ in the conjecture must be smaller than 1). To see that we can’t have $T \gg 2^\ell$, note that if we allow a NAND-CIRC program with much more than $2^\ell$ lines then this NAND-CIRC program could “hardwire” inside it all the outputs of $G$ on all its $2^\ell$ inputs, and use that to distinguish between a string of the form $G(s)$ and a uniformly chosen string in $\{0, 1\}^m$. To see that we can’t have $\epsilon \ll 2^{-\ell}$, note that by guessing the input $s$ (which will be successful with probability $2^{-2\ell}$), we can obtain a small (i.e., $O(\ell)$ line) NAND-CIRC program that achieves a $2^{-\ell}$ advantage in distinguishing a pseudorandom and uniform input. Working out these details is a highly recommended exercise.

We emphasize again that the optimal PRG conjecture is, as its name implies, a conjecture, and we still do not know how to prove it. In particular, it is stronger than the conjecture that $P \neq NP$. But we do have some evidence for its truth. There is a spectrum of different types of pseudorandom generators, and there are weaker assumptions than the optimal PRG conjecture that suffice to prove that $BPP = P$. In particular this is known to hold under the assumption that there exists a function $F \in \text{TIME}(2^{O(n)})$ and $\epsilon > 0$ such that for every sufficiently large $n$, $F_n$ is not in $\text{SIZE}(2^{\epsilon n})$. The name “Optimal PRG conjecture” is non standard. This conjecture is sometimes known in the literature as the existence of exponentially strong pseudorandom functions.\footnote{A pseudorandom generator of the form we posit, where each output bit can be computed individually in time polynomial in the seed length, is commonly known as a pseudorandom function generator. For more on the many interesting results and connections in the study of pseudorandomness, see this monograph of Salil Vadhan.}

19.3.3 Usefulness of pseudorandom generators

We now show that optimal pseudorandom generators are indeed very useful, by proving the following theorem:

**Theorem 19.10 — Derandomization of BPP.** Suppose that the optimal PRG conjecture is true. Then $BPP = P$.

**Proof Idea:**

The optimal PRG conjecture tells us that we can achieve exponential expansion of $\ell$ truly random coins into as many as $2^{\delta \ell}$ “pseudorandom coins.” Looked at from the other direction, it allows us to reduce the need for randomness by taking an algorithm that uses $m$ coins and converting it into an algorithm that only uses $O(\log m)$ coins. Now an algorithm of the latter type by can be made fully deterministic by enu-
merating over all the $2^{O(\log m)}$ (which is polynomial in $m$) possibilities for its random choices.

We now proceed with the proof details.

**Proof of Theorem 19.10.** Let $F \in \text{BPP}$ and let $P$ be a NAND-TM program and $a, b, c, d$ constants such that for every $x \in \{0, 1\}^n$, $P(x)$ runs in at most $c \cdot n^d$ steps and $\Pr_{r \sim \{0, 1\}^m}[P(x; r) = F(x)] \geq 2/3$.

By “unrolling the loop” and hardwiring the input $x$, we can obtain for every input $x \in \{0, 1\}^n$ a NAND-CIRC program $Q_x$ of at most, say, $T = 10c \cdot n^d$ lines, that takes $m$ bits of input and such that $Q(r) = P(x; r)$.

Now suppose that $G : \{0, 1\}^\ell \rightarrow \{0, 1\}$ is a $(T, 0.1)$ pseudorandom generator. Then we could deterministically estimate the probability $p(x) = \Pr_{s \sim \{0, 1\}^\ell}[Q_x(r) = 1]$ up to 0.1 accuracy in time $O(T \cdot 2^\ell \cdot m \cdot \text{cost}(G))$ where $\text{cost}(G)$ is the time that it takes to compute a single output bit of $G$.

The reason is that we know that $\tilde{p}(x) = \Pr_{s \sim \{0, 1\}^\ell}[Q_x(G(s)) = 1]$ will give us such an estimate for $p(x)$, and we can compute the probability $\tilde{p}(x)$ by simply trying all $2^\ell$ possibilities for $s$. Now, under the optimal PRG conjecture we can set $T = 2^{\ell/\delta}$ or equivalently $\ell = \frac{1}{\delta} \log T$, and our total computation time is polynomial in $2^\ell = T^{1/\delta}$.

Since $T \leq 10c \cdot n^d$, this running time will be polynomial in $n$.

This completes the proof, since we are guaranteed that $\Pr_{r \sim \{0, 1\}^m}[Q_x(r) = F(x)] \geq 2/3$, and hence estimating the probability $p(x)$ to within 0.1 accuracy is sufficient to compute $F(x)$.

---

### 19.4 P = NP AND BPP VS P

Two computational complexity questions that we cannot settle are:

- **Is $P = NP$?** Where we believe the answer is negative.
- **Is $BPP = P$?** Where we believe the answer is positive.

However we can say that the “conventional wisdom” is correct on at least one of these questions. Namely, if we’re wrong on the first count, then we’ll be right on the second one:

**Theorem 19.11 — Sipser–Gács Theorem.** If $P = NP$ then $BPP = P$.

Before reading the proof, it is instructive to think why this result is not “obvious.” If $P = NP$ then given any randomized algorithm $A$ and input $x$,
we will be able to figure out in polynomial time if there is a string \( r \in \{0, 1\}^m \) of random coins for \( A \) such that \( A(xr) = 1 \). The problem is that even if \( \Pr_{r \in \{0, 1\}^m}[A(xr) = F(x)] \geq 0.9999 \), it can still be the case that even when \( F(x) = 0 \) there exists a string \( r \) such that \( A(xr) = 1 \). The proof is rather subtle. It is much more important that you understand the statement of the theorem than that you follow all the details of the proof.

**Proof Idea:**

The construction follows the “quantifier elimination” idea which we have seen in Theorem 15.6. We will show that for every \( F \in \text{BPP} \), we can reduce the question of some input \( x \) satisfies \( F(x) = 1 \) to the question of whether a formula of the form \( \exists u \in \{0, 1\}^m \forall v \in \{0, 1\}^m P(u, v) \) is true, where \( m, k \) are polynomial in the length of \( x \) and \( P \) is polynomial-time computable. By Theorem 15.6, if \( P = \text{NP} \) then we can decide in polynomial time whether such a formula is true or false.

The idea behind this construction is that using amplification we can obtain a randomized algorithm \( A \) for computing \( F \) using \( m \) coins such that for every \( x \in \{0, 1\}^n \), if \( F(x) = 0 \) then the set \( S \subseteq \{0, 1\}^m \) of coins that make \( A \) output 1 is extremely tiny, and if \( F(x) = 1 \) then it is very large. Now in the case \( F(x) = 1 \), one can show that this means that there exists a small number \( k \) of “shifts” \( s_0, \ldots, s_{k-1} \) such that the union of the sets \( S \oplus s_i \) (i.e., sets of the form \( \{s \oplus s_i \mid s \in S\} \)) covers \( \{0, 1\}^m \), while in the case \( F(x) = 0 \) this union will always be of size at most \( k|S| \) which is much smaller than \( 2^m \). We can express the condition that there exists \( s_0, \ldots, s_{k-1} \) such that \( \bigcup_{i \in \{k\}} (S \oplus s_i) = \{0, 1\}^m \) as a statement with a constant number of quantifiers.

* 

**Proof of Theorem 19.11.** Let \( F \in \text{BPP} \). Using Theorem 19.4, there exists a polynomial-time algorithm \( A \) such that for every \( x \in \{0, 1\}^n \), \( \Pr_{x \in \{0, 1\}^n}[A(xr) = F(x)] \geq 1 - 2^{-n} \) where \( m \) is polynomial in \( n \). In particular (since an exponential dominates a polynomial, and we can always assume \( m \) is sufficiently large), it holds that

\[
\Pr_{x \in \{0, 1\}^n}[A(xr) = F(x)] \geq 1 - \frac{1}{10m^2}. \tag{19.7}
\]

Let \( x \in \{0, 1\}^n \), and let \( S_x \subseteq \{0, 1\}^m \) be the set \( \{r \in \{0, 1\}^m : A(xr) = 1\} \). By our assumption, if \( F(x) = 0 \) then \( |S_x| \leq \frac{1}{10m^2} 2^m \) and if \( F(x) = 1 \) then \( |S_x| \geq (1 - \frac{1}{10m^2}) 2^m \).

For a set \( S \subseteq \{0, 1\}^m \) and a string \( s \in \{0, 1\}^m \), we define the set \( S \oplus s \) to be \( \{r \oplus s : r \in S\} \) where \( \oplus \) denotes the XOR operation. That is, \( S \oplus s \) is the set \( S \) “shifted” by \( s \). Note that \( |S \oplus s| = |S| \). (Please make sure that you see why this is true.)
The heart of the proof is the following two claims:

**CLAIM I:** For every subset \( S \subseteq \{0, 1\}^m \), if \( |S| \leq \frac{1}{1000m^2} \), then for every \( s_0, \ldots, s_{100m-1} \in \{0, 1\}^m \), \( \cup_{i \in [100m]} (S \oplus s_i) \subseteq \{0, 1\}^m \).

**CLAIM II:** For every subset \( S \subseteq \{0, 1\}^m \), if \( |S| \geq \frac{1}{2} 2^m \), then there exist a set of string \( s_0, \ldots, s_{100m-1} \) such that \( \cup_{i \in [100m]} (S \oplus s_i) = \{0, 1\}^m \).

CLAIM I and CLAIM II together imply the theorem. Indeed, they mean that under our assumptions, for every \( x \in \{0, 1\}^n \), \( F(x) = 1 \) if and only if

\[
\exists s_0, \ldots, s_{100m-1} \in \{0, 1\}^m \cup_{i \in [100m]} (S_x \oplus s_i) = \{0, 1\}^m
\]

where we can re-write as

\[
\exists s_0, \ldots, s_{100m-1} \in \{0, 1\}^m \forall w \in \{0, 1\}^m (w \in (S_x \oplus s_0) \lor w \in (S_x \oplus s_1) \lor \ldots \lor w \in (S_x \oplus s_{100m-1}))
\]

or equivalently

\[
\exists s_0, \ldots, s_{100m-1} \in \{0, 1\}^m \forall w \in \{0, 1\}^m (A(x(w \oplus s_0)) = 1 \lor A(x(w \oplus s_1)) = 1 \lor \ldots \lor A(x(w \oplus s_{100m-1})) = 1)
\]

which (since \( A \) is computable in polynomial time) is exactly the type of statement shown in Theorem 15.6 to be decidable in polynomial time if \( P = NP \).

We see that all that is left is to prove **CLAIM I** and **CLAIM II**.

**CLAIM I** follows immediately from the fact that

\[
\cup_{i \in [100m-1]} |S_x \oplus s_i| \leq \sum_{i=0}^{100m-1} |S_x \oplus s_i| = \sum_{i=0}^{100m-1} |S_x| = 100m|S_x|
\]

To prove **CLAIM II**, we will use a technique known as the probabilistic method (see the proof of Lemma 19.12 for a more extensive discussion). Note that this is a completely different use of probability than in the theorem statement, we just use the methods of probability to prove an existential statement.

**Proof of CLAIM II:** Let \( S \subseteq \{0, 1\}^m \) with \( |S| \geq 0.5 \cdot 2^m \) be as in the claim’s statement. Consider the following probabilistic experiment: we choose \( 100m \) random shifts \( s_0, \ldots, s_{100m-1} \) independently at random in \( \{0, 1\}^m \), and consider the event \( GOOD \) that \( \cup_{i \in [100m]} (S \oplus s_i) = \{0, 1\}^m \). To prove CLAIM II it is enough to show that \( \Pr[GOOD] > 0 \), since that means that in particular there must exist shifts \( s_0, \ldots, s_{100m-1} \) that satisfy this condition.

For every \( z \in \{0, 1\}^m \), define the event \( BAD_z \) to hold if \( z \notin \cup_{i \in [100m]} (S \oplus s_i) \). The event \( GOOD \) holds if \( BAD_z \) fails for every \( z \in \{0, 1\}^m \), and so our goal is to prove that \( \Pr[\cup_{z \in \{0, 1\}^m} BAD_z] < 1 \). By
the union bound, to show this, it is enough to show that \( \Pr[BAD_z] < 2^{-m} \) for every \( z \in \{0, 1\}^m \). Define the event \( BAD_i^z \) to hold if \( z \notin S \oplus s_i \).

Since every shift \( s_i \) is chosen independently, for every fixed \( z \) the events \( BAD_z^0, \ldots, BAD_z^{100m-1} \) are mutually independent, and hence

\[
\Pr[BAD_z] = \Pr[\cap_{i=0}^{100m-1} BAD_i^z] = \prod_{i=0}^{100m-1} \Pr[BAD_i^z].
\]  

(19.12)

So this means that the result will follow by showing that \( \Pr[BAD_i^z] \leq \frac{1}{2} \) for every \( z \in \{0, 1\}^m \) and \( i \in [100m] \) (as that would allow to bound the righthand side of (19.12) by \( 2^{-100m} \)). In other words, we need to show that for every \( z \in \{0, 1\}^m \) and set \( S \subseteq \{0, 1\}^m \) with \( |S| \geq \frac{1}{2} 2^m \),

\[
\Pr_{s \in \{0,1\}^m} [z \in S \oplus s] \geq \frac{1}{2}.
\]  

(19.13)

To show this, we observe that \( z \in S \oplus s \) if and only if \( s \in S \oplus z \) (can you see why). Hence we can rewrite the probability on the lefthand side of (19.13) as \( \Pr_{s \in \{0,1\}^m} [s \in S \oplus z] \) which simply equals \( |S \oplus z|/2^m = |S|/2^m \geq 1/2 \). This concludes the proof of CLAIM I and hence of Theorem 19.11.

19.5 NON-CONSTRUCTIVE EXISTENCE OF PSEUDORANDOM GENERATORS (ADVANCED, OPTIONAL)

We now show that, if we don’t insist on constructivity of pseudorandom generators, then we can show that there exists pseudorandom generators with output that exponentially larger in the input length.

**Lemma 19.12 — Existence of inefficient pseudorandom generators.** There is some absolute constant \( C \) such that for every \( \epsilon, T \), if \( \ell > C(\log T + \log(1/\epsilon)) \) and \( m \leq T \), then there is an \( (T, \epsilon) \) pseudorandom generator \( G : \{0, 1\}^\ell \rightarrow \{0, 1\}^m \).

**Proof Idea:**

The proof uses an extremely useful technique known as the “probabilistic method” which is not too hard mathematically but can be confusing at first. The idea is to give a “non constructive” proof of existence of the pseudorandom generator \( G \) by showing that if \( G \) was chosen at random, then the probability that it would be a valid \((T, \epsilon) \) pseudorandom generator is positive. In particular this means that there exists a single \( G \) that is a valid \((T, \epsilon) \) pseudorandom generator. The probabilistic method is just a proof technique to demonstrate the existence of such a function. Ultimately, our goal is to show the existence of a deterministic function \( G \) that satisfies the condition.

\(^7\) There is a whole (highly recommended) book by Alon and Spencer devoted to this method.
The above discussion might be rather abstract at this point, but would become clearer after seeing the proof.

**Proof of Lemma 19.12.** Let $\epsilon, T, \ell, m$ be as in the lemma’s statement. We need to show that there exists a function $G : \{0, 1\}^\ell \to \{0, 1\}^m$ that “fools” every $T$ line program $P$ in the sense of (19.6). We will show that this follows from the following claim:

**Claim I:** For every fixed NAND-CIRC program $P$, if we pick $G : \{0, 1\}^\ell \to \{0, 1\}^m$ at random then the probability that (19.6) is violated is at most $2 - T^2$.

Before proving Claim I, let us see why it implies Lemma 19.12. We can identify a function $G : \{0, 1\}^\ell \to \{0, 1\}^m$ with its “truth table” or simply the list of evaluations on all its possible $2^\ell$ inputs. Since each output is an $m$ bit string, we can also think of $G$ as a string in $\{0, 1\}^{m \cdot 2^\ell}$. We define $\mathcal{F}_m^\ell$ to be the set of all functions from $\{0, 1\}^\ell$ to $\{0, 1\}^m$. As discussed above we can identify $\mathcal{F}_m^\ell$ with $\{0, 1\}^{m \cdot 2^\ell}$ and choosing a random function $G \sim \mathcal{F}_m^\ell$ corresponds to choosing a random $m \cdot 2^\ell$-long bit string.

For every NAND-CIRC program $P$ let $B_P$ be the event that, if we choose $G$ at random from $\mathcal{F}_m^\ell$ then (19.6) is violated with respect to the program $P$. It is important to understand what is the sample space that the event $B_P$ is defined over, namely this event depends on the choice of $G$ and so $B_P$ is a subset of $\mathcal{F}_m^\ell$. An equivalent way to define the event $B_P$ is that it is the subset of all functions mapping $\{0, 1\}^\ell$ to $\{0, 1\}^m$ that violate (19.6), or in other words:

$$B_P = \left\{ G \in \mathcal{F}_m^\ell \mid \frac{1}{2^\ell} \sum_{s \in \{0, 1\}^\ell} P(G(s)) - \frac{1}{2^m} \sum_{r \in \{0, 1\}^m} P(r) > \epsilon \right\}$$

(We’ve replaced here the probability statements in (19.6) with the equivalent sums so as to reduce confusion as to what is the sample space that $B_P$ is defined over.)

To understand this proof it is crucial that you pause here and see how the definition of $B_P$ above corresponds to (19.14). This may well take re-reading the above text once or twice, but it is a good exercise at parsing probabilistic statements and learning how to identify the sample space that these statements correspond to.

Now, we’ve shown in Theorem 5.2 that up to renaming variables (which makes no difference to program’s functionality) there are $2^{O(T \log T)}$ NAND-CIRC programs of at most $T$ lines. Since $T \log T < T^2$ for sufficiently large $T$, this means that if Claim I is true, then by the union bound it holds that the probability of the union of
$B_p$ over all NAND-CIRC programs of at most $T$ lines is at most $2^{O(T \log T)}2^{-T^2} < 0.1$ for sufficiently large $T$. What is important for us about the number 0.1 is that it is smaller than 1. In particular this means that there exists a single $G^* \in \mathcal{F}_m^T$ such that $G^*$ does not violate (19.6) with respect to any NAND-CIRC program of at most $T$ lines, but that precisely means that $G^*$ is a $(T, \epsilon)$ pseudorandom generator.

Hence to conclude the proof of Lemma 19.12, it suffices to prove Claim I. Choosing a random $G : \{0, 1\}^\ell \rightarrow \{0, 1\}^m$ amounts to choosing $L = 2^\ell$ random strings $y_0, \ldots, y_{L-1} \in \{0, 1\}^m$ and letting $G(x) = y_x$ (identifying $\{0, 1\}^\ell$ and $[L]$ via the binary representation). This means that proving the claim amounts to showing that for every fixed function $P : \{0, 1\}^m \rightarrow \{0, 1\}$, if $L > 2^{C(\log T + \log \epsilon)}$ (which by setting $C > 4$, we can ensure is larger than $10T^2/\epsilon^2$) then the probability that

$$\left| \frac{1}{L} \sum_{i=0}^{L-1} P(y_i) - \Pr_{s \sim \{0, 1\}^m}[P(s) = 1] \right| > \epsilon$$

is at most $2^{-T^2}$.

(??) follows directly from the Chernoff bound. Indeed, if we let for every $i \in [L]$ the random variable $X_i$ denote $P(y_i)$, then since $y_0, \ldots, y_{L-1}$ is chosen independently at random, these are independently and identically distributed random variables with mean $\mathbb{E}_{y \sim \{0, 1\}^m}[P(y)] = \Pr_{y \sim \{0, 1\}^m}[P(y) = 1]$ and hence the probability that they deviate from their expectation by $\epsilon$ is at most $2 \cdot 2^{-L/2}$.

\[\square\]

**Lecture Recap**

- We can model randomized algorithms by either adding a special “coin toss” operation or assuming an extra randomly chosen input.
- The class $\text{BPP}$ contains the set of Boolean functions that can be computed by polynomial time randomized algorithms.
- We know that $P \subseteq \text{BPP} \subseteq \text{EXP}$.
- We also know that $\text{BPP} \subseteq P_{/\text{poly}}$.
- The relation between $\text{BPP}$ and $\text{NP}$ is not known, but we do know that if $P = \text{NP}$ then $\text{BPP} = P$.
- Pseudorandom generators are objects that take a short random “seed” and expand it to a much longer output that “appears random” for efficient algorithms. We conjecture that exponentially strong pseudorandom generators exist. Under this conjecture, $\text{BPP} = P$. 

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**Modeling Randomized Computation**

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19.6 Exercises

19.7 Bibliographical Notes

19.8 Further Explorations

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

19.9 Acknowledgements
V

ADVANCED TOPICS
Cryptography - the art or science of “secret writing” - has been around for several millennia, and for almost all of that time Edgar Allan Poe’s quote above held true. Indeed, the history of cryptography is littered with the figurative corpses of cryptosystems believed secure and then broken, and sometimes with the actual corpses of those who have mistakenly placed their faith in these cryptosystems.

Yet, something changed in the last few decades, which is the “revolution” alluded to (and to a large extent initiated by) Diffie and Hellman’s 1976 paper quoted above. New cryptosystems have been found
that have not been broken despite being subjected to immense efforts involving both human ingenuity and computational power on a scale that completely dwarves the “code breakers” of Poe’s time. Even more amazingly, these cryptosystem are not only seemingly unbreakable, but they also achieve this under much harsher conditions. Not only do today’s attackers have more computational power but they also have more data to work with. In Poe’s age, an attacker would be lucky if they got access to more than a few encryptions of known messages. These days attackers might have massive amounts of data- terabytes or more - at their disposal. In fact, with public key encryption, an attacker can generate as many ciphertexts as they wish.

The key to this success has been a clearer understanding of both how to define security for cryptographic tools and how to relate this security to concrete computational problems. Cryptography is a vast and continuously changing topic, but we will touch on some of these issues in this chapter.

20.1 CLASSICAL CRYPTOSYSTEMS

A great many cryptosystems have been devised and broken throughout the ages. Let us recount just some of these stories. In 1587, Mary the queen of Scots, and the heir to the throne of England, wanted to arrange the assassination of her cousin, queen Elisabeth I of England, so that she could ascend to the throne and finally escape the house arrest under which she had been for the last 18 years. As part of this complicated plot, she sent a coded letter to Sir Anthony Babington. Mary used what’s known as a substitution cipher where each letter is transformed into a different obscure symbol (see Fig. 20.1). At a first look, such a letter might seem rather inscrutable- a meaningless sequence of strange symbols. However, after some thought, one might recognize that these symbols repeat several times and moreover that different symbols repeat with different frequencies. Now it doesn’t take a large leap of faith to assume that perhaps each symbol corresponds to a different letter and the more frequent symbols correspond to letters that occur in the alphabet with higher frequency. From this observation, there is a short gap to completely breaking the cipher, which was in fact done by queen Elisabeth’s spies who used the decoded letters to learn of all the co-conspirators and to convict queen Mary of treason, a crime for which she was executed. Trusting in superficial security measures (such as using “inscrutable” symbols) is a trap that users of cryptography have been falling into again and again over the years. (As in many things, this is the subject of a great XKCD cartoon, see Fig. 20.2.)

The Vigenère cipher is named after Blaise de Vigenère who described it in a book in 1586 (though it was invented earlier by Bellaso).
The idea is to use a collection of substitution cyphers - if there are $n$ different ciphers then the first letter of the plaintext is encoded with the first cipher, the second with the second cipher, the $n^{th}$ with the $n^{th}$ cipher, and then the $n + 1^{st}$ letter is again encoded with the first cipher. The key is usually a word or a phrase of $n$ letters, and the $i^{th}$ substitution cipher is obtained by shifting each letter $k_i$ positions in the alphabet. This “flattens” the frequencies and makes it much harder to do frequency analysis, which is why this cipher was considered “unbreakable” for 300+ years and got the nickname “le chiffre indéchiffrable” (“the unbreakable cipher”). Nevertheless, Charles Babbage cracked the Vigenère cipher in 1854 (though he did not publish it). In 1863 Friedrich Kasiski broke the cipher and published the result. The idea is that once you guess the length of the cipher, you can reduce the task to breaking a simple substitution cipher which can be done via frequency analysis (can you see why?). Confederate generals used Vigenère regularly during the civil war, and their messages were routinely cryptanalyzed by Union officers.

The Enigma cipher was a mechanical cipher (looking like a typewriter, see Fig. 20.5) where each letter typed would get mapped into a different letter depending on the (rather complicated) key and current state of the machine which had several rotors that rotated at different paces. An identically wired machine at the other end could be used to decrypt. Just as many ciphers in history, this has also been believed by the Germans to be “impossible to break” and even quite late in the war they refused to believe it was broken despite mounting evidence to that effect. (In fact, some German generals refused to believe it was broken even after the war.) Breaking Enigma was an heroic effort which was initiated by the Poles and then completed by the British at
Bletchley Park, with Alan Turing (of the Turing machines) playing a key role. As part of this effort the Brits built arguably the world’s first large scale mechanical computation devices (though they looked more similar to washing machines than to iPhones). They were also helped along the way by some quirks and errors of the German operators. For example, the fact that their messages ended with “Heil Hitler” turned out to be quite useful.

Here is one entertaining anecdote: the Enigma machine would never map a letter to itself. In March 1941, Mavis Batey, a cryptanalyst at Bletchley Park received a very long message that she tried to decrypt. She then noticed a curious property—the message did not contain the letter “L”\(^1\). She realized that the probability that no “L”’s appeared in the message is too small for this to happen by chance. Hence she surmised that the original message must have been composed only of L’s. That is, it must have been the case that the operator, perhaps to test the machine, have simply sent out a message where he repeatedly pressed the letter “L”. This observation helped her decode the next message, which helped inform of a planned Italian attack and secure a resounding British victory in what became known as “the Battle of Cape Matapan”. Mavis also helped break another Enigma machine. Using the information she provided, the Brits were able to feed the Germans with the false information that the main allied invasion would take place in Pas de Calais rather than on Normandy.

In the words of General Eisenhower, the intelligence from Bletchley park was of “priceless value”. It made a huge difference for the Allied war effort, thereby shortening World War II and saving millions of lives. See also this interview with Sir Harry Hinsley.

\(^1\) Here is a nice exercise: compute (up to an order of magnitude) the probability that a 50-letter long message composed of random letters will end up not containing the letter “L”.

**Figure 20.4**: Confederate encryption of the message “Gen’l Pemberton: You can expect no help from this side of the river. Let Gen’l Johnston know, if possible, when you can attack the same point on the enemy’s lines. Inform me also and I will endeavor to make a diversion. I have sent some caps. I subjoin a despatch from General Johnston.”

**Figure 20.5**: In the Enigma mechanical cipher the secret key would be the settings of the rotors and internal wires. As the operator types up their message, the encrypted version appeared in the display area above, and the internal state of the cipher was updated (and so typing the same letter twice would generally result in two different letters output). Decrypting follows the same process: if the sender and receiver are using the same key then typing the ciphertext would result in the plaintext appearing in the display.
20.2 DEFINING ENCRYPTION

Many of the troubles that cryptosystem designers faced over history (and still face!) can be attributed to not properly defining or understanding what are the goals they want to achieve in the first place. Let us focus on the setting of private key encryption. A sender (traditionally called “Alice”) wants to send a message (known also as a plaintext) $x \in \{0,1\}^*$ to a receiver (traditionally called “Bob”). They would like their message to be kept secret from an adversary who listens in or “eavesdrops” on the communication channel (and is traditionally called “Eve”).

Alice and Bob share a secret key $k \in \{0,1\}^*$. Alice uses the key $k$ to “scramble” or encrypt the plaintext $x$ into a ciphertext $y$, and Bob uses the key $k$ to “unscramble” or decrypt the ciphertext $y$ back into the plaintext $x$. This motivates the following definition which attempts to capture what it means for an encryption scheme to be valid or “make sense”, regardless of whether or not it is secure:

**Definition 20.1 — Valid encryption scheme.** Let $L : \mathbb{N} \to \mathbb{N}$ be some function. A pair of polynomial-time computable functions $(E, D)$ mapping strings to strings is a valid private key encryption scheme (or encryption scheme for short) with plaintext length function $L(\cdot)$ if for every $k \in \{0,1\}^n$ and $x \in \{0,1\}^{L(n)}$,

$$D(k, E(k, x)) = x.$$  

(20.1)

We also require that our encryption schemes are ciphertext length regular in the sense that all ciphertexts corresponding to keys of the same length are of the same length: there is some function $C : \mathbb{N} \to \mathbb{N}$ such that for every $k \in \{0,1\}^n$ and $x \in \{0,1\}^{L(n)}$,

$$|E(k, x)| = C(n).$$

We will often write the first input (i.e., the key) to the encryption and decryption as a subscript and so can write (20.1) also as $D_k(E_k(x)) = x$.

20.3 DEFINING SECURITY OF ENCRYPTION

Definition 20.1 says nothing about the security of $E$ and $D$, and even allows the trivial encryption scheme that ignores the key altogether and sets $E_k(x) = x$ for every $x$. Defining security is not a trivial matter.

2 If you don’t know what “private key” means, you can ignore this adjective for now. For thousands of years, “private key encryption” was synonymous with encryption. Only in the 1970’s was the concept of public key encryption invented, see Definition 20.10.

3 We call the function $L : \mathbb{N} \to \mathbb{N}$ the length function of $(E, D)$. The “ciphertext length regularity” condition is added for technical convenience and is not at all important. You can ignore it in a first reading.
Throughout history, many attacks on cryptosystems were rooted in the cryptosystem designers’ reliance on “security through obscurity”— trusting that the fact their methods are not known to their enemy will protect them from being broken. This is a faulty assumption - if you reuse a method again and again (even with a different key each time) then eventually your adversaries will figure out what you are doing. And if Alice and Bob meet frequently in a secure location to decide on a new method, they might as well take the opportunity to exchange their secrets. These considerations led Auguste Kerckhoffs in 1883 to state the following principle:

> A cryptosystem should be secure even if everything about the system, except the key, is public knowledge.  

Why is it OK to assume the key is secret and not the algorithm? Because we can always choose a fresh key. But of course that won’t help us much if our key is “1234” or “passw0rd!”. In fact, if you use any deterministic algorithm to choose the key then eventually your adversary will figure this out. Therefore for security we must choose the key at random and can restate Kerckhoffs’s principle as follows:

> There is no secrecy without randomness

This is such a crucial point that is worth repeating:

> There is no secrecy without randomness

At the heart of every cryptographic scheme there is a secret key, and the secret key is always chosen at random. A corollary of that is that to understand cryptography, you need to know probability theory.

**Remark 20.2 — Randomness in the real world.** Choosing the secrets for cryptography requires generating randomness, which is often done by measuring some “unpredictable” or “high entropy” data, and then applying hash functions to the result to “extract” a uniformly random string. Great care must be taken in doing this, and randomness generators often turn out to be the Achilles heel of secure systems. In 2006 a programmer removed a line of code from the procedure to generate entropy in OpenSSL package distributed by Debian since it caused a warning in

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4 The actual quote is “Il faut qu’il n’exige pas le secret, et qu’il puisse sans inconvénient tomber entre les mains de l’ennemi” loosely translated as “The system must not require secrecy and can be stolen by the enemy without causing trouble”. According to Steve Bellovin the NSA version is “assume that the first copy of any device we make is shipped to the Kremlin”.

some automatic verification code. As a result for two years (until this was discovered) all the randomness generated by this procedure used only the process ID as an “unpredictable” source. This means that all communication done by users in that period is fairly easily breakable (and in particular, if some entities recorded that communication they could break it also retroactively). See XKCD’s take on that incident.

In 2012 two separate teams of researchers scanned a large number of RSA keys on the web and found out that about 4 percent of them are easy to break. The main issue were devices such as routers, internet-connected printers and such. These devices sometimes run variants of Linux-a desktop operating system—but without a hard drive, mouse or keyboard, they don’t have access to many of the entropy sources that desktop have. Coupled with some good old fashioned ignorance of cryptography and software bugs, this led to many keys that are downright trivial to break, see this blog post and this web page for more details.

Since randomness is so crucial to security, breaking the procedure to generate randomness can lead to a complete break of the system that uses this randomness. Indeed, the Snowden documents, combined with observations of Shumow and Ferguson, strongly suggest that the NSA has deliberately inserted a trapdoor in one of the pseudorandom generators published by the National Institute of Standards and Technologies (NIST). Fortunately, this generator wasn’t widely adapted but apparently the NSA did pay 10 million dollars to RSA security so the latter would make this generator their default option in their products.

20.4 PERFECT SECRECY

If you think about encryption scheme security for a while, you might come up with the following principle for defining security: “An encryption scheme is secure if it is not possible to recover the key \( k \) from \( E_k(x) \).” However, a moment’s thought shows that the key is not really what we’re trying to protect. After all, the whole point of an encryption is to protect the confidentiality of the plaintext \( x \). So, we can try to define that “an encryption scheme is secure if it is not possible to recover the plaintext \( x \) from \( E_k(x) \).” Yet it is not clear what this means either. Suppose that an encryption scheme reveals the first 10 bits of the plaintext \( x \). It might still not be possible to recover \( x \) completely, but on an intuitive level, this seems like it would be extremely unwise to use such an encryption scheme in practice. Indeed, often even partial information about the plaintext is enough for the adversary to achieve its goals.

The above thinking led Shannon in 1945 to formalize the notion of perfect secrecy, which is that an encryption reveals absolutely nothing
about the message. There are several equivalent ways to define it, but perhaps the cleanest one is the following:

**Definition 20.3 — Perfect secrecy.** A valid encryption scheme \((E, D)\) with length \(L(\cdot)\) is perfectly secret if for every \(n \in \mathbb{N}\) and plaintexts \(x, x' \in \{0, 1\}^{L(n)}\), the following two distributions \(Y\) and \(Y'\) over \(\{0, 1\}^*\) are identical:

- \(Y\) is obtained by sampling \(k \sim \{0, 1\}^n\) and outputting \(E_k(x)\).
- \(Y'\) is obtained by sampling \(k \sim \{0, 1\}^n\) and outputting \(E_k(x')\).

This definition might take more than one reading to parse. Try to think of how this condition would correspond to your intuitive notion of “learning no information” about \(x\) from observing \(E_k(x)\), and to Shannon’s quote in the beginning of this chapter. In particular, suppose that you knew ahead of time that Alice sent either an encryption of \(x\) or an encryption of \(x'\). Would you learn anything new from observing the encryption of the message that Alice actually sent? It may help you to look at Fig. 20.6.

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**20.4.1 Example: Perfect secrecy in the battlefield**

To understand Definition 20.3, suppose that Alice sends only one of two possible messages: “attack” or “retreat”, which we denote by \(x_0\) and \(x_1\) respectively, and that she sends each one of those messages with probability \(1/2\). Let us put ourselves in the shoes of Eve, the eavesdropping adversary. A priori we would have guessed that Alice sent either \(x_0\) or \(x_1\) with probability \(1/2\). Now we observe \(y = E_k(x_i)\) where \(k\) is a uniformly chosen key in \(\{0, 1\}^n\). How does this new information cause us to update our beliefs on whether Alice sent the plaintext \(x_0\) or the plaintext \(x_1\)?

Let us define \(p_0(y)\) to be the probability (taken over \(k \sim \{0, 1\}^n\)) that \(y = E_k(x_0)\) and similarly \(p_1(y)\) to be \(\text{Pr}_{k \sim \{0, 1\}^n}[y = E_k(x_1)]\). Note that, since Alice chooses the message to send at random, our a priori probability for observing \(y\) is \(\frac{1}{2}p_0(y) + \frac{1}{2}p_1(y)\). However,
as per Definition 20.3, the perfect secrecy condition guarantees that $p_0(y) = p_1(y)$! Let us denote the number $p_0(y) = p_1(y)$ by $p$. By the formula for conditional probability, the probability that Alice sent the message $x_0$ conditioned on our observation $y$ is simply\(^5\)

$$\Pr[i = 0|y = E_k(x_i)] = \frac{\Pr[i = 0 \land y = E_k(x_i)]}{\Pr[y = E_k(x)]}. \quad (20.2)$$

Since the probability that $i = 0$ and $y$ is the ciphertext $E_k(0)$ is equal to $\frac{1}{2} \cdot p_0(y)$, and the a priori probability of observing $y$ is $\frac{1}{2} p_0(y) + \frac{1}{2} p_1(y)$, we can rewrite (20.2) as

$$\Pr[i = 0|y = E_k(x_i)] = \frac{\frac{1}{2}p_0(y)}{\frac{1}{2}p_0(y) + \frac{1}{2}p_1(y)} = \frac{p}{p + p} = \frac{1}{2} \quad (20.3)$$

using the fact that $p_0(y) = p_1(y) = p$. This means that observing the ciphertext $y$ did not help us at all! We still would not be able to guess whether Alice sent “attack” or “retreat” with better than 50/50 odds!

This example can be vastly generalized to show that perfect secrecy is indeed “perfect” in the sense that observing a ciphertext gives Eve no additional information about the plaintext beyond her a priori knowledge.

20.4.2 Constructing perfectly secret encryption

Perfect secrecy is an extremely strong condition, and implies that an eavesdropper does not learn any information from observing the ciphertext. You might think that an encryption scheme satisfying such a strong condition will be impossible, or at least extremely complicated, to achieve. However it turns out we can in fact obtain perfectly secret encryption scheme fairly easily. Such a scheme for two-bit messages is illustrated in Fig. 20.7.

In fact, this can be generalized to any number of bits:

\begin{quote}
**Theorem 20.4 — One Time Pad (Vernam 1917, Shannon 1949).** There is a perfectly secret valid encryption scheme $(E, D)$ with $L(n) = n$.
\end{quote}

**Proof Idea:**

Our scheme is the one-time pad also known as the “Vernam Cipher”, see Fig. 20.8. The encryption is exceedingly simple: to encrypt a message $x \in \{0, 1\}^n$ with a key $k \in \{0, 1\}^n$ we simply output $x \oplus k$ where $\oplus$ is the bitwise XOR operation that outputs the string corresponding to XORing each coordinate of $x$ and $k$.

Proof of Theorem 20.4. For two binary strings $a$ and $b$ of the same length $n$, we define $a \oplus b$ to be the string $c \in \{0, 1\}^n$ such that

\begin{align*}
\Pr[i = 0|y = E_k(x_i)] &= \frac{\frac{1}{2}p_0(y)}{\frac{1}{2}p_0(y) + \frac{1}{2}p_1(y)} = \frac{p}{p + p} = \frac{1}{2} \quad (20.3)
\end{align*}

\begin{align*}
\text{The equation (20.2) is a special case of Bayes’ rule which, although a simple restatement of the formula for conditional probability, is an extremely important and widely used tool in statistics and data analysis.}
\end{align*}
The argument above is quite simple but is worth reading again. To understand why the one-time pad is perfectly secret, it is useful to envision it as a bipartite graph as we’ve done in Fig. 20.7. (In fact the encryption scheme of Fig. 20.7 is precisely the one-time pad for $n = 2$.) For every $n$, the one-time pad encryption scheme corresponds to a bipartite graph with $2^n$ vertices on the “left side” corresponding to the plaintexts in ${0, 1}^n$ and $2^n$ vertices on the “right side” corresponding to the ciphertexts ${0, 1}^n$. For every $x \in {0, 1}^n$ and $k \in {0, 1}^n$, we connect $x$ to the vertex $y = E_k(x)$ with an edge that we label with $k$. One can see that this is the complete bipartite graph, where every vertex on the left is connected to all vertices on the right. In particular this means that for every left vertex $x$, the distribution on the ciphertexts obtained by taking a random $k \in {0, 1}^n$ and going to the neighbor of $x$ on the edge labeled $k$ is the uniform distribution over ${0, 1}^n$. This ensures the perfect secrecy condition.

20.5 NECESSITY OF LONG KEYS

So, does Theorem 20.4 give the final word on cryptography, and means that we can all communicate with perfect secrecy and live happily ever after? No it doesn’t. While the one-time pad is efficient, and gives perfect secrecy, it has one glaring disadvantage: to communicate $n$ bits you need to store a key of length $n$. In contrast, practically used cryptosystems such as AES-128 have a short key of 128 bits (i.e.,
16 bytes) that can be used to protect terabytes or more of communication! Imagine that we all needed to use the one time pad. If that was the case, then if you had to communicate with \( m \) people, you would have to maintain (securely!) \( m \) huge files that are each as long as the length of the maximum total communication you expect with that person. Imagine that every time you opened an account with Amazon, Google, or any other service, they would need to send you in the mail (ideally with a secure courier) a DVD full of random numbers, and every time you suspected a virus, you’d need to ask all these services for a fresh DVD. This doesn’t sound so appealing.

This is not just a theoretical issue. The Soviets have used the one-time pad for their confidential communication since before the 1940’s. In fact, even before Shannon’s work, the U.S. intelligence already knew in 1941 that the one-time pad is in principle “unbreakable” (see page 32 in the Venona document). However, it turned out that the hassle of manufacturing so many keys for all the communication took its toll on the Soviets and they ended up reusing the same keys for more than one message. They did try to use them for completely different receivers in the (false) hope that this wouldn’t be detected. The Venona Project of the U.S. Army was founded in February 1943 by Gene Grabeel (see Fig. 20.9), a former home economics teacher from Madison Heights, Virginia and Lt. Leonard Zubko. In October 1943, they had their breakthrough when it was discovered that the Russians were reusing their keys.\(^6\) In the 37 years of its existence, the project has resulted in a treasure chest of intelligence, exposing hundreds of KGB agents and Russian spies in the U.S. and other countries, including Julius Rosenberg, Harry Gold, Klaus Fuchs, Alger Hiss, Harry Dexter White and many others.

Unfortunately it turns out that that such long keys are necessary for perfect secrecy:

\[ L(n) \leq n. \]

**Proof Idea:**

The idea behind the proof is illustrated in Fig. 20.10. If the number of keys is smaller than the number of messages then the neighborhoods of all vertices in the corresponding graphs cannot be identical.

\[ \star \]

**Proof of Theorem 20.5.** Let \( E, D \) be a valid encryption scheme with messages of length \( L \) and key of length \( n < L \). We will show that \( (E, D) \) is not perfectly secret by providing two plaintexts \( x_0, x_1 \in \]
\{0, 1\}^L$ such that the distributions $Y_{x_0}$ and $Y_{x_1}$ are not identical, where $Y_x$ is the distribution obtained by picking $k \sim \{0, 1\}^n$ and outputting $E_k(x)$.

We choose $x_0 = 0^L$. Let $S_0 \subseteq \{0, 1\}^*$ be the set of all ciphertexts that have nonzero probability of being output in $Y_{x_0}$. That is, $S_0 = \{y \mid \exists k \in \{0, 1\}^n : y = E_k(x_0)\}$. Since there are only $2^n$ keys, we know that $|S_0| \leq 2^n$.

We will show the following claim:

**Claim I:** There exists some $x_1 \in \{0, 1\}^L$ and $k \in \{0, 1\}^n$ such that $E_k(x_1) \notin S_0$.

Claim I implies that the string $E_k(x_1)$ has positive probability of being output by $Y_{x_1}$ and zero probability of being output by $Y_{x_0}$ and hence in particular $Y_{x_0}$ and $Y_{x_1}$ are not identical. To prove Claim I, just choose a fixed $k \in \{0, 1\}^n$. By the validity condition, the map $x \mapsto E_k(x)$ is a one to one map of $\{0, 1\}^L$ to $\{0, 1\}^*$ and hence in particular the image of this map which is the set $I_k = \{y \mid \exists x \in \{0, 1\}^L : y = E_k(x)\}$ has size at least (in fact exactly) $2^L$. Since $|S_0| \leq 2^n < 2^L$, this means that $|I_k| > |S_0|$ and so in particular there exists some string $y$ in $I_k \setminus S_0$. But by the definition of $I_k$ this means that there is some $x \in \{0, 1\}^L$ such that $E_k(x) \notin S_0$ which concludes the proof of Claim I and hence of Theorem 20.5.

20.6 **COMPUTATIONAL SECRECY**

To sum up the previous episodes, we now know that:

- It is possible to obtain a perfectly secret encryption scheme with key length the same as the plaintext.

- It is not possible to obtain such a scheme with key that is even a single bit shorter than the plaintext.

How does this mesh with the fact that, as we’ve already seen, people routinely use cryptosystems with a 16 byte (i.e., 128 bit) key but many terabytes of plaintext? The proof of Theorem 20.5 does give in fact a way to break all these cryptosystems, but an examination of this proof shows that it only yields an algorithm with time exponential in the length of the key. This motivates the following relaxation of perfect secrecy to a condition known as “computational secrecy”. Intuitively, an encryption scheme is computationally secret if no polynomial time algorithm can break it. The formal definition is below:
**Definition 20.6 — Computational secrecy.** Let \((E, D)\) be a valid encryption scheme where for keys of length \(n\), the plaintexts are of length \(L(n)\) and the ciphertexts are of length \(m(n)\). We say that \((E, D)\) is **computationally secret** if for every polynomial \(p : \mathbb{N} \to \mathbb{N}\), and large enough \(n\), if \(P\) is an \(m(n)\)-input and single output NAND-CIRC program of at most \(p(n)\) lines, and \(x_0, x_1 \in \{0, 1\}^{L(n)}\) then

\[
\left| \mathbb{E}_{k \sim \{0,1\}^n} [P(E_k(x_0))] - \mathbb{E}_{k \sim \{0,1\}^n} [P(E_k(x_1))] \right| < \frac{1}{p(n)} \quad (20.4)
\]

**P**

Definition 20.6 requires a second or third read and some practice to truly understand. One excellent exercise to make sure you follow it is to see that if we allow \(P\) to be an arbitrary function mapping \(\{0, 1\}^{m(n)}\) to \(\{0, 1\}\), and we replace the condition in (20.4) that the lefthand side is smaller than \(\frac{1}{p(L(n))}\) with the condition that it is equal to 0 then we get the perfect secrecy condition of Definition 20.3. Indeed if the distributions \(E_k(x_0)\) and \(E_k(x_1)\) are identical then applying any function \(P\) to them we get the same expectation. On the other hand, if the two distributions above give a different probability for some element \(y^* \in \{0, 1\}^{m(n)}\), then the function \(P(y)\) that outputs 1 iff \(y = y^*\) will have a different expectation under the former distribution than under the latter.

**Definition 20.6** raises two natural questions:

- Is it strong enough to ensure that a computationally secret encryption scheme protects the secrecy of messages that are encrypted with it?

- It is weak enough that, unlike perfect secrecy, it is possible to obtain a computationally secret encryption scheme where the key is much smaller than the message?

To the best of our knowledge, the answer to both questions is **Yes**. Regarding the first question, it is not hard to show that if, for example, Alice uses a computationally secret encryption algorithm to encrypt either “attack” or “retreat” (each chosen with probability 1/2), then as long as she’s restricted to polynomial-time algorithms, an adversary Eve will not be able to guess the message with probability better than, say, 0.51, even after observing its encrypted form. (We omit the proof, but it is an excellent exercise for you to work it out on your own.)
To answer the second question we will show that under the same assumption we used for derandomizing BPP, we can obtain a computationally secret cryptosystem where the key is almost exponentially smaller than the plaintext.

20.6.1 Stream ciphers or the “derandomized one-time pad”

It turns out that if pseudorandom generators exist as in the optimal PRG conjecture, then there exists a computationally secret encryption scheme with keys that are much shorter than the plaintext. The construction below is known as a stream cipher, though perhaps a better name is the “derandomized one-time pad”. It is widely used in practice with keys on the order of a few tens or hundreds of bits protecting many terabytes or even petabytes of communication.

Theorem 20.7 — Derandomized one-time pad. Suppose that the optimal PRG conjecture is true. Then for every constant \( a \in \mathbb{N} \) there is a computationally secret encryption scheme \((E, D)\) with plaintext length \( L(n) \) at least \( n^a \).

Proof Idea:
The proof is illustrated in Fig. 20.11. We simply take the one-time pad on \( L \) bit plaintexts, but replace the key with \( G(k) \) where \( k \) is a string in \( \{0, 1\}^n \) and \( G : \{0, 1\}^n \to \{0, 1\}^L \) is a pseudorandom generator. Since the one time pad cannot be broken, an adversary that breaks the derandomized one-time pad can be used to distinguish between the output of the pseudorandom generator and the uniform distribution.

* 

Proof of Theorem 20.7. Since an exponential function of the form \( 2^{\delta n} \) grows faster than any polynomial of the form \( n^a \), under the optimal PRG conjecture we can obtain a polynomial-time computable \((2^{\delta n}, 2^{-\delta n})\) pseudorandom generator \( G : \{0, 1\}^n \to \{0, 1\}^L \) for \( L = n^a \). We now define our encryption scheme as follows: given key \( k \in \{0, 1\}^n \) and plaintext \( x \in \{0, 1\}^L \), the encryption \( E_k(x) \) is simply \( x \oplus G(k) \). To decrypt a string \( y \in \{0, 1\}^m \) we output \( y \oplus G(k) \).

This is a valid encryption since \( G \) is computable in polynomial time and \( (x \oplus G(k)) \oplus G(k) = x \oplus (G(k) \oplus G(k)) = x \) for every \( x \in \{0, 1\}^L \).

Computational secrecy follows from the condition of a pseudorandom generator. Suppose, towards a contradiction, that there is a polynomial \( p \), NAND-CIRC program \( Q \) of at most \( p(L) \) lines and \( x, x' \in \{0, 1\}^{L(n)} \) such that

\[
\left| \mathbb{E}_{k \sim \{0, 1\}^n} [Q(E_k(x))] - \mathbb{E}_{k \sim \{0, 1\}^n} [Q(E_k(x'))] \right| > \frac{1}{p(L)}.
\] (20.5)
(We use here the simple fact that for a \( \{0, 1\} \)-valued random variable \( X \), \( \Pr[X = 1] = \mathbb{E}[X] \).)

By the definition of our encryption scheme, this means that

\[
\left| \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x)] - \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x')] \right| > \frac{1}{p(L)}. \tag{20.6}
\]

Now since (as we saw in the security analysis of the one-time pad), for every strings \( x, x' \in \{0, 1\}^L \), the distribution \( r \oplus x \) and \( r \oplus x' \) are identical, where \( r \sim \{0, 1\}^L \). Hence

\[
\mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x)] = \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x')]. \tag{20.7}
\]

By plugging (20.7) into (20.6) we can derive that

\[
\left| \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x)] - \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x)] + \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x')] - \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x')] \right| > \frac{1}{p(L)}. \tag{20.8}
\]

(Please make sure that you can see why this is true.)

Now we can use the \textit{triangle inequality} that \( |A + B| \leq |A| + |B| \) for every two numbers \( A, B \), applying it for \( A = \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x)] - \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x)] \) and \( B = \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x')] - \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x')] \)

to derive

\[
\mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x)] - \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x)] + \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x')] - \mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x')] > \frac{1}{p(L)}. \tag{20.9}
\]

In particular, either the first term or the second term of the lefthand-side of (20.9) must be at least \( \frac{1}{2p(L)} \). Let us assume the first case holds (the second case is analyzed in exactly the same way).

Then we get that

\[
\mathbb{E}_{k \sim \{0,1\}^n} [Q(G(k) \oplus x)] - \mathbb{E}_{r \sim \{0,1\}^L} [Q(r \oplus x)] > \frac{1}{2p(L)}. \tag{20.10}
\]

But if we now define the NAND-CIRC program \( P_x \) that on input \( r \in \{0, 1\}^L \) outputs \( Q(r \oplus x) \) then (since XOR of \( L \) bits can be computed in \( O(L) \) lines), we get that \( P_x \) has \( p(L) + O(L) \) lines and by (20.10) it can distinguish between an input of the form \( G(k) \) and an input of the form \( r \sim \{0, 1\}^k \) with advantage better than \( \frac{1}{2p(L)} \). Since a polynomial is dominated by an exponential, if we make \( L \) large enough, this will contradict the \((2^{dn}, 2^{-dn})\) security of the pseudorandom generator \( G \).}

\( \blacksquare \)

**Remark 20.8 — Stream ciphers in practice.** The two most widely used forms of (private key) encryption schemes in practice are \textit{stream ciphers} and \textit{block ciphers}.
We model breaking the scheme as Eve outputting 0 or 1 corresponding to whether the message sent was $x_0$ or $x_1$. Note that we could have just as well modified Eve to output $x_0$ instead of 0 and $x_1$ instead of 1. The key point is that a priori Eve only had a 50/50 chance of guessing whether Alice sent $x_0$ or $x_1$ but after seeing the ciphertext this chance increases to better than 99/1.

7 We model breaking the scheme as Eve outputting 0 or 1 corresponding to whether the message sent was $x_0$ or $x_1$. Note that we could have just as well modified Eve to output $x_0$ instead of 0 and $x_1$ instead of 1. The key point is that a priori Eve only had a 50/50 chance of guessing whether Alice sent $x_0$ or $x_1$ but after seeing the ciphertext this chance increases to better than 99/1.

20.7 COMPUTATIONAL SECRECY AND NP

We’ve also mentioned before that an efficient algorithm for NP could be used to break all cryptography. We now give an example of how this can be done:

**Theorem 20.9 — Breaking encryption using NP algorithm.** If $P = NP$ then there is no computationally secret encryption scheme with $L(n) > n$.

Furthermore, for every valid encryption scheme $(E, D)$ with $L(n) > n + 100$ there is a polynomial $p$ such that for every large enough $n$ there exist $x_0, x_1 \in \{0, 1\}^{L(n)}$ and a $p(n)$-line NAND-CIRC program $EVE$ s.t.

$$\Pr_{i \sim \{0, 1\}, k \sim \{0, 1\}^n} \left[ EVE(E_k(x_i)) = i \right] \geq 0.99 .$$

(20.11)

Note that the “furthermore” part is extremely strong. It means that if the plaintext is even a little bit larger than the key, then we can already break the scheme in a very strong way. That is, there will be a pair of messages $x_0, x_1$ (think of $x_0$ as “sell” and $x_1$ as “buy”) and an efficient strategy for Eve such that if Eve gets a ciphertext $y$ then she will be able to tell whether $y$ is an encryption of $x_0$ or $x_1$ with probability very close to 1. The condition $P = NP$ can be relaxed to $NP \subseteq BPP$ and even the weaker condition $NP \subseteq P/poly$ with essentially the same proof.

**Proof Idea:**
The proof follows along the lines of Theorem 20.5 but this time paying attention to the computational aspects. If $P = \text{NP}$ then for every plaintext $x$ and ciphertext $y$, we can efficiently tell whether there exists $k \in \{0, 1\}^n$ such that $E_k(x) = y$. So, to prove this result we need to show that if the plaintexts are long enough, there would exist a pair $x_0, x_1$ such that the probability that a random encryption of $x_1$ also is a valid encryption of $x_0$ will be very small. The details of how to show this are below.

* 

Proof of Theorem 20.9. We focus on showing only the “furthermore” part since it is the more interesting and the other part follows by essentially the same proof.

Suppose that $(E, D)$ is such an encryption, let $n$ be large enough, and let $x_0 = 0^{L(n)}$. For every $x \in \{0, 1\}^{L(n)}$ we define $S_x$ to be the set of all valid encryption of $x$. That is $S_x = \{y \mid \exists k \in \{0, 1\}^n \ y = E_k(x)\}$. As in the proof of Theorem 20.5, since there are $2^n$ keys $k$, $|S_x| \leq 2^n$ for every $x \in \{0, 1\}^{L(n)}$.

We denote by $S_0$ the set $S_{x_0}$. We define our algorithm $\text{EVE}$ to output 0 on input $y \in \{0, 1\}^*$ if $y \in S_0$ and to output 1 otherwise. This can be implemented in polynomial time if $P = \text{NP}$, since the key $k$ can serve the role of an efficiently verifiable solution. (Can you see why?) Clearly $\Pr[\text{EVE}(E_k(x_0)) = 0] = 1$ and so in the case that $\text{EVE}$ gets an encryption of $x_0$ then she guesses correctly with probability 1. The remainder of the proof is devoted to showing that there exists $x_1 \in \{0, 1\}^{L(n)}$ such that $\Pr[\text{EVE}(E_k(x_1)) = 0] \leq 0.01$, which will conclude the proof by showing that $\text{EVE}$ guesses wrongly with probability at most $\frac{1}{2}0 + \frac{1}{2}0.01 < 0.01$.

Consider now the following probabilistic experiment (which we define solely for the sake of analysis). We consider the sample space of choosing $x$ uniformly in $\{0, 1\}^{L(n)}$ and define the random variable $Z_k(x)$ to equal 1 if and only if $E_k(x) \in S_0$. For every $k$, the map $x \mapsto E_k(x)$ is one-to-one, which means that the probability that $Z_k = 1$ is equal to the probability that $x \in E_k^{-1}(S_0)$ which is $|S_0|^{2^n}$. So by the linearity of expectation $\mathbb{E}\left[\sum_{k \in \{0, 1\}^n} Z_k\right] \leq 2^n |S_0|^{2^n} \leq \frac{2^n}{2^{2n}}$.

We will now use the following extremely simple but useful fact known as the averaging principle (see also ??): for every random variable $Z$, if $\mathbb{E}[Z] = \mu$, then with positive probability $Z \leq \mu$. (Indeed, if $Z > \mu$ with probability one, then the expected value of $Z$ will have to be larger than $\mu$, just like you can’t have a class in which all students got A or A- and yet the overall average is B+.) In our case it means that with positive probability $\sum_{k \in \{0, 1\}^n} Z_k \leq \frac{2^n}{2^{2n}}$. In other words, there exists some $x_1 \in \{0, 1\}^{L(n)}$ such that $\sum_{k \in \{0, 1\}^n} Z_k(x_1) \leq \frac{2^n}{2^{2n}}$. Yet this means that if we choose a random $k \sim \{0, 1\}^n$, then the
probability that $E_k(x_1) \in S_0$ is at most $\frac{1}{2} \cdot \frac{2^{2n}}{2^{L(n)}} = 2^{n-L(n)}$. So, in particular if we have an algorithm $EVE$ that outputs 0 if $x \in S_0$ and outputs 1 otherwise, then $\Pr[EVE(E_k(x) = 0] = 1$ and $\Pr[EVE(E_k(x_1)) = 0] \leq 2^{n-L(n)}$ which will be smaller than $2^{-10} < 0.01$ if $L(n) \geq n+10$.

In retrospect Theorem 20.9 is perhaps not surprising. After all, as we’ve mentioned before it is known that the Optimal PRG conjecture (which is the basis for the derandomized one-time pad encryption) is false if $P = NP$ (and in fact even if $NP \subseteq BPP$ or even $NP \subseteq P/poly$).

20.8 PUBLIC KEY CRYPTOGRAPHY

People have been dreaming about heavier-than-air flight since at least the days of Leonardo Da Vinci (not to mention Icarus from the greek mythology). Jules Verne wrote with rather insightful details about going to the moon in 1865. But, as far as I know, in all the thousands of years people have been using secret writing, until about 50 years ago no one has considered the possibility of communicating securely without first exchanging a shared secret key.

Yet in the late 1960’s and early 1970’s, several people started to question this “common wisdom”. Perhaps the most surprising of these visionaries was an undergraduate student at Berkeley named Ralph Merkle. In the fall of 1974 Merkle wrote in a project proposal for his computer security course that while “it might seem intuitively obvious that if two people have never had the opportunity to prearrange an encryption method, then they will be unable to communicate securely over an insecure channel... I believe it is false”. The project proposal was rejected by his professor as “not good enough”. Merkle later submitted a paper to the communication of the ACM where he apologized for the lack of references since he was unable to find any mention of the problem in the scientific literature, and the only source where he saw the problem even raised was in a science fiction story. The paper was rejected with the comment that “Experience shows that it is extremely dangerous to transmit key information in the clear.”

Merkle showed that one can design a protocol where Alice and Bob can use $T$ invocations of a hash function to exchange a key, but an adversary (in the random oracle model, though he of course didn’t use this name) would need roughly $T^2$ invocations to break it. He conjectured that it may be possible to obtain such protocols where breaking is exponentially harder than using them, but could not think of any concrete way to doing so.

We only found out much later that in the late 1960’s, a few years before Merkle, James Ellis of the British Intelligence agency GCHQ
was having similar thoughts. His curiosity was spurred by an old World-War II manuscript from Bell labs that suggested the following way that two people could communicate securely over a phone line. Alice would inject noise to the line, Bob would relay his messages, and then Alice would subtract the noise to get the signal. The idea is that an adversary over the line sees only the sum of Alice’s and Bob’s signals, and doesn’t know what came from what. This got James Ellis thinking whether it would be possible to achieve something like that digitally. As Ellis later recollected, in 1970 he realized that in principle this should be possible, since he could think of an hypothetical black box $B$ that on input a “handle” $\alpha$ and plaintext $x$ would give a “ciphertext” $y$ and that there would be a secret key $\beta$ corresponding to $\alpha$, such that feeding $\beta$ and $y$ to the box would recover $x$. However, Ellis had no idea how to actually instantiate this box. He and others kept giving this question as a puzzle to bright new recruits until one of them, Clifford Cocks, came up in 1973 with a candidate solution loosely based on the factoring problem; in 1974 another GCHQ recruit, Malcolm Williamson, came up with a solution using modular exponentiation.

But among all those thinking of public key cryptography, probably the people who saw the furthest were two researchers at Stanford, Whit Diffie and Martin Hellman. They realized that with the advent of electronic communication, cryptography would find new applications beyond the military domain of spies and submarines, and they understood that in this new world of many users and point to point communication, cryptography will need to scale up. Diffie and Hellman envisioned an object which we now call “trapdoor permutation” though they called “one way trapdoor function” or sometimes simply “public key encryption”. Though they didn’t have full formal definitions, their idea was that this is an injective function that is easy (e.g., polynomial-time) to compute but hard (e.g., exponential-time) to invert. However, there is a certain trapdoor, knowledge of which would allow polynomial time inversion. Diffie and Hellman argued that using such a trapdoor function, it would be possible for Alice and Bob to communicate securely without ever having exchanged a secret key. But they didn’t stop there. They realized that protecting the integrity of communication is no less important than protecting its secrecy. Thus they imagined that Alice could “run encryption in reverse” in order to certify or sign messages.

At the point, Diffie and Hellman were in a position not unlike physicists who predicted that a certain particle should exist but without any experimental verification. Luckily they met Ralph Merkle, and his ideas about a probabilistic key exchange protocol, together with a suggestion from their Stanford colleague John Gill, inspired them
to come up with what today is known as the Diffie Hellman Key Exchange (which unbeknownst to them was found two years earlier at GCHQ by Malcolm Williamson). They published their paper “New Directions in Cryptography” in 1976, and it is considered to have brought about the birth of modern cryptography.

The Diffie-Hellman Key Exchange is still widely used today for secure communication. However, it still felt short of providing Diffie and Hellman’s elusive trapdoor function. This was done the next year by Rivest, Shamir and Adleman who came up with the RSA trapdoor function, which through the framework of Diffie and Hellman yielded not just encryption but also signatures. From this point on began a flurry of advances in cryptography which hasn’t died down till this day.

20.8.1 Defining public key encryption

A public key encryption consists of a triple of algorithms:

- The key generation algorithm, which we denote by KeyGen or KG for short, is a randomized algorithm that outputs a pair of strings \((e, d)\) where \(e\) is known as the public (or encryption) key, and \(d\) is known as the private (or decryption) key. The key generation algorithm gets as input \(1^n\) (i.e., a string of ones of length \(n\)). We refer to \(n\) as the security parameter of the scheme. The bigger we make \(n\), the more secure the encryption will be, but also the less efficient it will be.

- The encryption algorithm, which we denote by \(E\), takes the encryption key \(e\) and a plaintext \(x\), and outputs the ciphertext \(y = E_e(x)\).

- The decryption algorithm, which we denote by \(D\), takes the decryption key \(d\) and a ciphertext \(y\), and outputs the plaintext \(x = D_d(y)\).

We now make this a formal definition:

**Definition 20.10 — Public Key Encryption.** A computationally secret public key encryption with plaintext length \(L : \mathbb{N} \to \mathbb{N}\) is a triple of randomized polynomial-time algorithms \((KG, E, D)\) that satisfy the following conditions:

- For every \(n\), if \((e, d)\) is output by \(KG(1^n)\) with positive probability, and \(x \in \{0, 1\}^{L(n)}\), then \(D_d(E_e(x)) = x\) with probability one.

- For every polynomial \(p\), and sufficiently large \(n\), if \(P\) is a NAND-CIRC program of at most \(p(n)\) lines then for every \(x, x' \in \{0, 1\}^{L(n)}\), \(D_d(E_e(x)) = x\) with probability one.
If you want to learn more about the different types of public key assumptions, you can take a look at my own survey on this topic.

One way to compute a logarithm is by binary search: start with some interval $[x_{\text{min}}, x_{\text{max}}]$ that is guaranteed to contain $\log_y g$. We can then test whether the interval’s midpoint $x_{\text{mid}}$ satisfies $g^{x_{\text{mid}}} > y$, and based on that halve the size of the interval.

This is not trivial, and is a great exercise for you to work this out. As a hint, start by showing that one can compute the map $k \mapsto g^{2^k} \mod p$ using $k$ modular multiplications modulo $p$. If you’re stumped, you can look up this Wikipedia entry.

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**Definition 20.10** allows $E$ and $D$ to be randomized algorithms. In fact, it turns out that it is necessary for $E$ to be randomized to obtain computational secrecy. It also turns out that, unlike the private key case, we can transform a public-key encryption that works for messages that are only one bit long into a public-key encryption scheme that can encrypt arbitrarily long messages, and in particular messages that are longer than the key. In particular this means that we cannot obtain a perfectly secret public-key encryption scheme even for one-bit long messages (since it would imply a perfectly secret public-key, and hence in particular private-key, encryption with messages longer than the key).

We will not give full constructions for public key encryption schemes in this chapter, but will mention some of the ideas that underlie the most widely used schemes today. These generally belong to one of two families:

- **Group theoretic constructions** based on problems such as integer factoring and the discrete logarithm over finite fields or elliptic curves.
- **Lattice/coding based constructions** based on problems such as the closest vector in a lattice or bounded distance decoding.

Group-theory based encryptions such as the RSA cryptosystem, the Diffie-Hellman protocol, and Elliptic-Curve Cryptography, are currently more widely implemented. But the lattice/coding schemes are recently on the rise, particularly because the known group theoretic encryption schemes can be broken by quantum computers, which we’ll discuss later in this course.

### 20.8.2 Diffie-Hellman key exchange

As just one example of how public key encryption schemes are constructed, let us now describe the Diffie-Hellman key exchange. We describe the Diffie-Hellman protocol in a somewhat of an informal level, without presenting a full security analysis.

The computational problem underlying the Diffie-Hellman protocol is the discrete logarithm problem. Let’s suppose that $g$ is some integer. We can compute the map $x \mapsto g^x$ and also its inverse $y \mapsto \log_g y$. However, suppose now that we use modular arithmetic and work modulo some prime number $p$. If $p$ has $n$ binary digits and $g$ is in $[p]$ then we can compute the map $x \mapsto g^x \mod p$ in time polynomial in $n$.

On the other hand, because of the “wraparound” property of modular arithmetic, we cannot run binary search to find the inverse of

$$\{0, 1\}^{L(n)}, |E[P(e, E_e(x))] - E[P(e, E'_e(x'))]| < 1/p(n),$$

where this probability is taken over the coins of $KG$ and $E$.  

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9 If you want to learn more about the different types of public key assumptions, you can take a look at my own survey on this topic.

10 One way to compute a logarithm is by binary search: start with some interval $[x_{\text{min}}, x_{\text{max}}]$ that is guaranteed to contain $\log_y g$. We can then test whether the interval’s midpoint $x_{\text{mid}}$ satisfies $g^{x_{\text{mid}}} > y$, and based on that halve the size of the interval.

11 This is not trivial, and is a great exercise for you to work this out. As a hint, start by showing that one can compute the map $k \mapsto g^{2^k} \mod p$ using $k$ modular multiplications modulo $p$. If you’re stumped, you can look up this Wikipedia entry.
this map (known as the discrete logarithm). In fact, there is no known polynomial-time algorithm for computing this discrete logarithm map \( (g, x, p) \mapsto \log_g x \mod p \), where we define \( \log_g x \mod p \) as the number \( a \in [p] \) such that \( g^a = x \mod p \).

The Diffie-Hellman protocol for Bob to send a message to Alice is as follows:

- **Alice:** Chooses \( p \) to be a random \( n \) bit long prime (which can be done by choosing random numbers and running a primality testing algorithm on them), and \( g \) and \( a \) at random in \([p]\). She sends to Bob the triple \((p, g, g^a \mod p)\).

- **Bob:** Given the triple \((p, g, h)\), Bob sends a message \( x \in \{0, 1\}^L \) to Alice by choosing \( b \) at random in \([p]\), and sending to Alice the pair \((g^b \mod p, rep(h^b \mod p) \oplus x)\) where \( rep : [p] \to \{0, 1\}^* \) is some “representation function” that maps \([p]\) to \( \{0, 1\}^L \). 12

- **Alice:** Given \( g', z \), Alice recovers \( x \) by outputting \( rep(g'^a \mod p) \oplus z \).

The correctness of the protocol follows from the simple fact that \((g^a)^b = (g^b)^a\) for every \( g, a, b \) and this still holds if we work modulo \( p \). Its security relies on the computational assumption that computing this map is hard, even in a certain “average case” sense (this computational assumption is known as the Decisional Diffie Hellman assumption). The Diffie-Hellman key exchange protocol can be thought of as a public key encryption where the Alice’s first message is the public key, and Bob’s message is the encryption.

One can think of the Diffie-Hellman protocol as being based on a “trapdoor pseudorandom generator” whereas the triple \( g^a, g^b, g^{ab} \) looks “random” to someone that doesn’t know \( a \), but someone that does know \( a \) can see that raising the second element to the \( a \)-th power yields the third element. The Diffie-Hellman protocol can be described abstractly in the context of any finite Abelian group for which we can efficiently compute the group operation. It has been implemented on other groups than numbers modulo \( p \), and in particular Elliptic Curve Cryptography (ECC) is obtained by basing the Diffie Hellman on elliptic curve groups which gives some practical advantages. 13

Another common group theoretic basis for key-exchange/public key encryption protocol is the RSA function. A big disadvantage of Diffie-Hellman (both the modular arithmetic and elliptic curve variants) and RSA is that both schemes can be broken in polynomial time by a quantum computer. We will discuss quantum computing later in this course.

12 The function \( rep \) does not need to be one-to-one and you can think of \( rep(z) \) as simply outputting \( L \) of the bits of \( z \) in the natural binary representation. The function \( rep \) does need to satisfy certain technical conditions which we omit in this description.

13 The main advantage in ECC is that the best known algorithms for computing discrete logarithms over elliptic curve groups take time \( 2^{\epsilon n} \) for some \( \epsilon > 0 \) where \( n \) is the number of bits to describe a group element. In contrast, for the multiplicative group modulo a prime \( p \) the best algorithm take time \( O(n^{1/3} \text{poly}(\log(n))) \) which means that (assuming the known algorithms are optimal) we need to set the prime to be bigger (and so have larger key sizes with corresponding overhead in communication and computation) to get the same level of security.
20.9 OTHER SECURITY NOTIONS

There is a great deal to cryptography beyond just encryption schemes, and beyond the notion of a passive adversary. A central objective is integrity or authentication: protecting communications from being modified by an adversary. Integrity is often more fundamental than secrecy: whether it is a software update or viewing the news, you might often not care about the communication being secret as much as that it indeed came from its claimed source. Digital signature schemes are the analog of public key encryption for authentication, and are widely used (in particular as the basis for public key certificates) to provide a foundation of trust in the digital world.

Similarly, even for encryption, we often need to ensure security against active attacks, and so notions such as non-malleability and adaptive chosen ciphertext security have been proposed. An encryption scheme is only as secure as the secret key, and mechanisms to make sure the key is generated properly, and is protected against refresh or even compromise (i.e., forward secrecy) have been studied as well. Hopefully this chapter provides you with some appreciation for cryptography as an intellectual field, but does not imbue you with a false self of confidence in implementing it.

Cryptographic hash functions is another widely used tool with a variety of uses, including extracting randomness from high entropy sources, achieving hard-to-forge short “digests” of files, protecting passwords, and much more.

20.10 MAGIC

Beyond encryption and signature schemes, cryptographers have managed to obtain objects that truly seem paradoxical and “magical”. We briefly discuss some of these objects. We do not give any details, but hopefully this will spark your curiosity to find out more.

20.10.1 Zero knowledge proofs

On October 31, 1903, the mathematician Frank Nelson Cole, gave an hourlong lecture to a meeting of the American Mathematical Society where he did not speak a single word. Rather, he calculated on the board the value $2^{67} - 1$ which is equal to 147,573,952,589,676,412,927, and then showed that this number is equal to $193,707,721 \times 761,838,257,287$. Cole’s proof showed that $2^{67} - 1$ is not a prime, but it also revealed additional information, namely its actual factors. This is often the case with proofs: they teach us more than just the validity of the statements.

In Zero Knowledge Proofs we try to achieve the opposite effect. We want a proof for a statement $X$ where we can rigorously show that the
proofs reveals absolutely no additional information about $X$ beyond the fact that it is true. This turns out to be an extremely useful object for a variety of tasks including authentication, secure protocols, voting, anonymity in cryptocurrencies, and more. Constructing these objects relies on the theory of NP completeness. Thus this theory that originally was designed to give a negative result (show that some problems are hard) ended up yielding positive applications, enabling us to achieve tasks that were not possible otherwise.

20.10.2 Fully homomorphic encryption
Suppose that we are given a bit-by-bit encryption of a string $E_k(x_0), \ldots, E_k(x_{n-1})$. By design, these ciphertexts are supposed to be “completely unscrutable” and we should not be able to extract any information about $x_i$’s from it. However, already in 1978, Rivest, Adleman and Dertouzos observed that this does not imply that we could not manipulate these encryptions. For example, it turns out the security of an encryption scheme does not immediately rule out the ability to take a pair of encryptions $E_k(a)$ and $E_k(b)$ and compute from them $E_k(a \text{NAND} b)$ without knowing the secret key $k$. But do there exist encryption schemes that allow such manipulations? And if so, is this a bug or a feature?

Rivest et al already showed that such encryption schemes could be immensely useful, and their utility has only grown in the age of cloud computing. After all, if we can compute NAND then we can use this to run any algorithm $P$ on the encrypted data, and map $E_k(x_0), \ldots, E_k(x_{n-1})$ to $E_k(P(x_0, \ldots, x_{n-1}))$. For example, a client could store their secret data $x$ in encrypted form on the cloud, and have the cloud provider perform all sorts of computation on these data without ever revealing to the provider the private key, and so without the provider ever learning any information about the secret data.

The question of existence of such a scheme took much longer time to resolve. Only in 2009 Craig Gentry gave the first construction of an encryption scheme that allows to compute a universal basis of gates on the data (known as a Fully Homomorphic Encryption scheme in crypto parlance). Gentry’s scheme left much to be desired in terms of efficiency, and improving upon it has been the focus of an intensive research program that has already seen significant improvements.

20.10.3 Multiparty secure computation
Cryptography is about enabling mutually distrusting parties to achieve a common goal. Perhaps the most general primitive achieving this objective is secure multiparty computation. The idea in secure multiparty computation is that $n$ parties interact together to compute
some function $F(x_0, \ldots, x_{n-1})$ where $x_i$ is the private input of the $i$-th party. The crucial point is that there is no commonly trusted party or authority and that nothing is revealed about the secret data beyond the function’s output. One example is an electronic voting protocol where only the total vote count is revealed, with the privacy of the individual voters protected, but without having to trust any authority to either count the votes correctly or to keep information confidential. Another example is implementing a second price (aka Vickrey) auction where $n - 1$ parties submit bids to an item owned by the $n$-th party, and the item goes to the highest bidder but at the price of the second highest bid. Using secure multiparty computation we can implement second price auction in a way that will ensure the secrecy of the numerical values of all bids (including even the top one) except the second highest one, and the secrecy of the identity of all bidders (including even the second highest bidder) except the top one. We emphasize that such a protocol requires no trust even in the auctioneer itself, that will also not learn any additional information. Secure multiparty computation can be used even for computing randomized processes, with one example being playing Poker over the net without having to trust any server for correct shuffling of cards or not revealing the information.

### Lecture Recap

- We can formally define the notion of security of an encryption scheme.
- **Perfect secrecy** ensures that an adversary does not learn anything about the plaintext from the ciphertext, regardless of their computational powers.
- The one-time pad is a perfectly secret encryption with the length of the key equaling the length of the message. No perfectly secret encryption can have key shorter than the message.
- **Computational secrecy** can be as good as perfect secrecy since it ensures that the advantage that computationally bounded adversaries gain from observing the ciphertext is exponentially small. If the optimal PRG conjecture is true then there exists a computationally secret encryption scheme with messages that can be (almost) exponentially bigger than the key.
- There are many cryptographic tools that go well beyond private key encryption. These include public key encryption, digital signatures and hash functions, as well as more “magical” tools such as multiparty secure computation, fully homomorphic encryption, zero knowledge proofs, and many others.
Much of this text is taken from my lecture notes on cryptography. Shannon’s manuscript was written in 1945 but was classified, and a partial version was only published in 1949. Still it has revolutionized cryptography, and is the forerunner to much of what followed.

John Nash made seminal contributions in mathematics and game theory, and was awarded both the Abel Prize in mathematics and the Nobel Memorial Prize in Economic Sciences. However, he has struggled with mental illness throughout his life. His biography, A Beautiful Mind was made into a popular movie. It is natural to compare Nash’s 1955 letter to the NSA to Gödel’s letter to von Neumann we mentioned before. From the theoretical computer science point of view, the crucial difference is that while Nash informally talks about exponential vs polynomial computation time, he does not mention the word “Turing Machine” or other models of computation, and it is not clear if he is aware or not that his conjecture can be made mathematically precise (assuming a formalization of “sufficiently complex types of enciphering”).

The definition of computational secrecy we use is the notion of computational indistinguishability (known to be equivalent to semantic security) that was given by Goldwasser and Micali in 1982.

Although they used a different terminology, Diffie and Hellman already made clear in their paper that their protocol can be used as a public key encryption, with the first message being put in a “public file”. In 1985, ElGamal showed how to obtain a signature scheme based on the Diffie Hellman ideas, and since he described the Diffie-Hellman encryption scheme in the same paper, it is sometimes also known as ElGamal encryption.

Zero-knowledge proofs were constructed by Goldwasser, Micali, and Rackoff in 1982, and their wide applicability was shown (using the theory of NP completeness) by Goldreich, Micali, and Wigderson in 1986.

Two party and multiparty secure computation protocols were constructed (respectively) by Yao in 1982 and Goldreich, Micali, and Wigderson in 1987. The latter work gave a general transformation from security against passive adversaries to security against active adversaries using zero knowledge proofs.

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)
20.14 ACKNOWLEDGEMENTS
21
Proofs and algorithms

“Let’s not try to define knowledge, but try to define zero-knowledge.”, Shafi Goldwasser.

Proofs have captured human imagination for thousands of years, ever since the publication of Euclid’s Elements, a book second only to the bible in the number of editions printed.

Plan:

• Proofs and algorithms
• Interactive proofs
• Zero knowledge proofs
• Propositions as types, Coq and other proof assistants.

21.1 LECTURE SUMMARY
21.2 EXERCISES
21.3 BIBLIOGRAPHICAL NOTES
21.4 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

21.5 ACKNOWLEDGEMENTS
Quantum computing

“We always have had (secret, secret, close the doors!) ... a great deal of difficulty in understanding the world view that quantum mechanics represents ... It has not yet become obvious to me that there’s no real problem. ... Can I learn anything from asking this question about computers--about this may or may not be mystery as to what the world view of quantum mechanics is?”, Richard Feynman, 1981

“The only difference between a probabilistic classical world and the equations of the quantum world is that somehow or other it appears as if the probabilities would have to go negative”, Richard Feynman, 1981

There were two schools of natural philosophy in ancient Greece. Aristotle believed that objects have an essence that explains their behavior, and a theory of the natural world has to refer to the reasons (or “final cause” to use Aristotle’s language) as to why they exhibit certain phenomena. Democritus believed in a purely mechanistic explanation of the world. In his view, the universe was ultimately composed of elementary particles (or Atoms) and our observed phenomena arise from the interactions between these particles according to some local rules. Modern science (arguably starting with Newton) has embraced Democritus’ point of view, of a mechanistic or “clockwork” universe of particles and forces acting upon them.

While the classification of particles and forces evolved with time, to a large extent the “big picture” has not changed from Newton till Einstein. In particular it was held as an axiom that if we knew fully the current state of the universe (i.e., the particles and their properties such as location and velocity) then we could predict its future state at any point in time. In computational language, in all these theories the state of a system with \(n\) particles could be stored in an array of \(O(n)\) numbers, and predicting the evolution of the system can be done by

Learning Objectives:
- See main aspects in which quantum mechanics differs from local deterministic theories.
- Model of quantum circuits, or equivalently QNAND-CIRC programs
- The complexity class BQP and what we know about its relation to other classes
- Ideas behind Shor’s Algorithm and the Quantum Fourier Transform
running some efficient (e.g., $poly(n)$ time) deterministic computation on this array.

## 22.1 THE DOUBLE SLIT EXPERIMENT

Alas, in the beginning of the 20th century, several experimental results were calling into question this “clockwork” or “billiard ball” theory of the world. One such experiment is the famous double slit experiment. Here is one way to describe it. Suppose that we buy one of those baseball pitching machines, and aim it at a soft plastic wall, but put a metal barrier with a single slit between the machine and the plastic wall (see Fig. 22.1). If we shoot baseballs at the plastic wall, then some of the baseballs would bounce off the metal barrier, while some would make it through the slit and dent the wall. If we now carve out an additional slit in the metal barrier then more balls would get through, and so the plastic wall would be even more dented.

So far this is pure common sense, and it is indeed (to my knowledge) an accurate description of what happens when we shoot baseballs at a plastic wall. However, this is not the same when we shoot photons. Amazingly, if we shoot with a “photon gun” (i.e., a laser) at a wall equipped with photon detectors through some barrier, then (as shown in Fig. 22.2) in some positions of the wall we will see fewer hits when the two slits are open than when only one of them is! In particular there are positions in the wall that are hit when the first slit is open, hit when the second gun is open, but are not hit at all when both slits are open.

It seems as if each photon coming out of the gun is aware of the global setup of the experiment, and behaves differently if two slits are open than if only one is. If we try to “catch the photon in the act” and place a detector right next to each slit so we can see exactly the path each photon takes then something even more bizarre happens. The mere fact that we measure the path changes the photon’s behavior, and now this “destructive interference” pattern is gone and the number of times a position is hit when two slits are open is the sum of the number of times it is hit when each slit is open.

![Figure 22.1](image1.png)

Figure 22.1: In the “double baseball experiment” we shoot baseballs from a gun at a soft wall through a hard barrier that has one or two slits open in it. There is only “constructive interference” in the sense that the dent in each position in the wall when both slits are open is the sum of the dents when each slit is open on its own.

You should read the paragraphs above more than once and make sure you appreciate how truly mind boggling these results are.

## 22.2 QUANTUM AMPLITUDES

The double slit and other experiments ultimately forced scientists to accept a very counterintuitive picture of the world. It is not merely
about nature being randomized, but rather it is about the probabilities in some sense “going negative” and cancelling each other!

To see what we mean by this, let us go back to the baseball experiment. Suppose that the probability a ball passes through the left slit is \( p_L \) and the probability that it passes through the right slit is \( p_R \). Then, if we shoot \( N \) balls out of each gun, we expect the wall will be hit \((p_L + p_R)N\) times. In contrast, in the quantum world of photons instead of baseballs, it can sometimes be the case that in both the first and second case the wall is hit with positive probabilities \( p_L \) and \( p_R \) respectively but somehow when both slits are open the wall (or a particular position in it) is not hit at all. It’s almost as if the probabilities can “cancel each other out”.

To understand the way we model this in quantum mechanics, it is helpful to think of a “lazy evaluation” approach to probability. We can think of a probabilistic experiment such as shooting a baseball through two slits in two different ways:

- When a ball is shot, “nature” tosses a coin and decides if it will go through the left slit (which happens with probability \( p_L \)), right slit (which happens with probability \( p_R \)), or bounce back. If it passes through one of the slits then it will hit the wall. Later we can look at the wall and find out whether or not this event happened, but the fact that the event happened or not is determined independently of whether or not we look at the wall.

- The other viewpoint is that when a ball is shot, “nature” computes the probabilities \( p_L \) and \( p_R \) as before, but does not yet “toss the coin” and determines what happened. Only when we actually look at the wall, nature tosses a coin and with probability \( p_L + p_R \) ensures we see a dent. That is, nature uses “lazy evaluation”, and only determines the result of a probabilistic experiment when we decide to measure it.

While the first scenario seems much more natural, the end result in both is the same (the wall is hit with probability \( p_L + p_R \)) and so the question of whether we should model nature as following the first scenario or second one seems like asking about the proverbial tree that falls in the forest with no one hearing about it.

However, when we want to describe the double slit experiment with photons rather than baseballs, it is the second scenario that lends itself better to a quantum generalization. Quantum mechanics associates a number \( \alpha \) known as an amplitude with each probabilistic experiment. This number \( \alpha \) can be negative, and in fact even complex. We never observe the amplitudes directly, since whenever we measure an event with amplitude \( \alpha \), nature tosses a coin and determines
that the event happens with probability $|\alpha|^2$. However, the sign (or in the complex case, phase) of the amplitudes can affect whether two different events have constructive or destructive interference.

Specifically, consider an event that can either occur or not (e.g., “detector number 17 was hit by a photon”). In classical probability, we model this by a probability distribution over the two outcomes: a pair of non-negative numbers $p$ and $q$ such that $p + q = 1$, where $p$ corresponds to the probability that the event occurs and $q$ corresponds to the probability that the event does not occur. In quantum mechanics, we model this also by pair of numbers, which we call amplitudes. This is a pair of (potentially negative or even complex) numbers $\alpha$ and $\beta$ such that $|\alpha|^2 + |\beta|^2 = 1$. The probability that the event occurs is $|\alpha|^2$ and the probability that it does not occur is $|\beta|^2$. In isolation, these negative or complex numbers don’t matter much, since we anyway square them to obtain probabilities. But the interaction of positive and negative amplitudes can result in surprising cancellations where somehow combining two scenarios where an event happens with positive probability results in a scenario where it never does.

If you don’t find the above description confusing and unintuitive, you probably didn’t get it. Please make sure to re-read the above paragraphs until you are thoroughly confused.

Quantum mechanics is a mathematical theory that allows us to calculate and predict the results of the double-slit and many other experiments. If you think of quantum mechanics as an explanation as to what “really” goes on in the world, it can be rather confusing. However, if you simply “shut up and calculate” then it works amazingly well at predicting experimental results. In particular, in the double slit experiment, for any position in the wall, we can compute numbers $\alpha$ and $\beta$ such that photons from the first and second slit hit that position with probabilities $|\alpha|^2$ and $|\beta|^2$ respectively. When we open both slits, the probability that the position will be hit is proportional to $|\alpha + \beta|^2$, and so in particular, if $\alpha = -\beta$ then it will be the case that, despite being hit when either slit one or slit two are open, the position is not hit at all when they both are. If you are confused by quantum mechanics, you are not alone: for decades people have been trying to come up with explanations for “the underlying reality” behind quantum mechanics, including Bohmian Mechanics, Many Worlds and others. However, none of these interpretations have gained universal acceptance and all of those (by design) yield the same experimental predictions. Thus at this point many scientists prefer to just ignore the
question of what is the “true reality” and go back to simply “shutting up and calculating”.

Remark 22.1 — Complex vs real, other simplifications. If (like the author) you are a bit intimidated by complex numbers, don’t worry: you can think of all amplitudes as real (though potentially negative) numbers without loss of understanding. All the “magic” of quantum computing already arises in this case, and so we will often restrict attention to real amplitudes in this chapter.

We will also only discuss so-called pure quantum states, and not the more general notion of mixed states. Pure states turn out to be sufficient for understanding the algorithmic aspects of quantum computing. More generally, this chapter is not meant to be a complete description of quantum mechanics, quantum information theory, or quantum computing, but rather illustrate the main points where these differ from classical computing.

22.3 BELL’S INEQUALITY

There is something weird about quantum mechanics. In 1935 Einstein, Podolsky and Rosen (EPR) tried to pinpoint this issue by highlighting a previously unrealized corollary of this theory. They showed that the idea that nature does not determine the results of an experiment until it is measured results in so called “spooky action at a distance”. Namely, making a measurement of one object may instantaneously effect the state (i.e., the vector of amplitudes) of another object in the other end of the universe.

Since the vector of amplitudes is just a mathematical abstraction, the EPR paper was considered to be merely a thought experiment for philosophers to be concerned about, without bearing on experiments. This changed when in 1965 John Bell showed an actual experiment to test the predictions of EPR and hence pit intuitive common sense against quantum mechanics. Quantum mechanics won: it turns out that it is in fact possible to use measurements to create correlations between the states of objects far removed from one another that cannot be explained by any prior theory. Nonetheless, since the results of these experiments are so obviously wrong to anyone that has ever sat in an armchair, that there are still a number of Bell denialists arguing that this can’t be true and quantum mechanics is wrong.

So, what is this Bell’s Inequality? Suppose that Alice and Bob try to convince you they have telepathic ability, and they aim to prove it via the following experiment. Alice and Bob will be in separate closed
rooms. You will interrogate Alice and your associate will interrogate Bob. You choose a random bit \( x \in \{0, 1\} \) and your associate chooses a random \( y \in \{0, 1\} \). We let \( a \) be Alice’s response and \( b \) be Bob’s response. We say that Alice and Bob win this experiment if \( a \oplus b = x \land y \). In other words, Alice and Bob need to output two bits that disagree if \( x = y = 1 \) and agree otherwise.

Now if Alice and Bob are not telepathic, then they need to agree in advance on some strategy. It’s not hard for Alice and Bob to succeed with probability \( 3/4 \): just always output the same bit. Moreover, by doing some case analysis, we can show that no matter what strategy they use, Alice and Bob cannot succeed with higher probability than that.

**Theorem 22.2 — Bell’s Inequality.** For every two functions \( f, g : \{0, 1\} \to \{0, 1\} \), \( \Pr_{x,y \in \{0,1\}}[f(x) \oplus g(y) = x \land y] \leq 3/4 \).

**Proof.** Since the probability is taken over all four choices of \( x, y \in \{0, 1\} \), the only way the theorem can be violated if there exist two functions \( f, g \) that satisfy

\[
f(x) \oplus g(y) = x \land y
\]

for all the four choices of \( x, y \in \{0, 1\} \). Let’s plug in all these four choices and see what we get (below we use the equalities \( z \oplus 0 = z \), \( z \land 0 = 0 \) and \( z \land 1 = z \)):

\[
\begin{align*}
  f(0) \oplus g(0) &= 0 \quad \text{(plugging in } x = 0, y = 0) \\
  f(0) \oplus g(1) &= 0 \quad \text{(plugging in } x = 0, y = 1) \\
  f(1) \oplus g(0) &= 0 \quad \text{(plugging in } x = 1, y = 0) \\
  f(1) \oplus g(1) &= 1 \quad \text{(plugging in } x = 1, y = 1)
\end{align*}
\]

(22.1)

If we XOR together the first and second equalities we get \( g(0) \oplus g(1) = 0 \) while if we XOR together the third and fourth equalities we get \( g(0) \oplus g(1) = 1 \), thus obtaining a contradiction.

An amazing experimentally verified fact is that quantum mechanics allows for “telepathy.” Specifically, it has been shown that using the weirdness of quantum mechanics, there is in fact a strategy for Alice and Bob to succeed in this game with probability larger than \( 3/4 \) (in fact, they can succeed with probability about \( 0.85 \), see Lemma 22.5).

**22.4 QUANTUM WEIRDNESS**

Some of the counterintuitive properties that arise from quantum mechanics include:
• **Interference** - As we’ve seen, quantum amplitudes can “cancel each other out”.

• **Measurement** - The idea that amplitudes are negative as long as “no one is looking” and “collapse” (by squaring them) to positive probabilities when they are measured is deeply disturbing. Indeed, as shown by EPR and Bell, this leads to various strange outcomes such as “spooky actions at a distance”, where we can create correlations between the results of measurements in places far removed. Unfortunately (or fortunately?) these strange outcomes have been confirmed experimentally.

• **Entanglement** - The notion that two parts of the system could be connected in this weird way where measuring one will affect the other is known as *quantum entanglement*.

As counter-intuitive as these concepts are, they have been experimentally confirmed, so we just have to live with them.

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**Remark 22.3 — More on quantum.** The discussion in this lecture is quite brief and somewhat superficial. The chapter on quantum computation in my book with Arora (see draft here) is one relatively short resource that contains essentially everything we discuss here and more. See also this blog post of Aaronson for a high level explanation of Shor’s algorithm which ends with links to several more detailed expositions. This lecture of Aaronson contains a great discussion of the feasibility of quantum computing (Aaronson’s course lecture notes and the book that they spawned are fantastic reads as well). The videos of Umesh Vazirani’s EdX course are an accessible and recommended introduction to quantum computing. See the “bibliographical notes” section at the end of this chapter for more resources.

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**22.5 QUANTUM COMPUTING AND COMPUTATION - AN EXECUTIVE SUMMARY.**

One of the strange aspects of the quantum-mechanical picture of the world is that unlike in the billiard ball example, there is no obvious algorithm to simulate the evolution of $n$ particles over $t$ time periods in $\text{poly}(n, t)$ steps. In fact, the natural way to simulate $n$ quantum particles will require a number of steps that is exponential in $n$. This is a huge headache for scientists that actually need to do these calculations in practice.
In the 1981, physicist Richard Feynman proposed to “turn this lemon to lemonade” by making the following almost tautological observation:

If a physical system cannot be simulated by a computer in \( T \) steps, the system can be considered as performing a computation that would take more than \( T \) steps.

So, he asked whether one could design a quantum system such that its outcome \( y \) based on the initial condition \( x \) would be some function \( y = f(x) \) such that (a) we don’t know how to efficiently compute in any other way, and (b) is actually useful for something. In 1985, David Deutsch formally suggested the notion of a quantum Turing machine, and the model has been since refined in works of Deutsch, Josza, Bernstein and Vazirani. Such a system is now known as a quantum computer.

For a while these hypothetical quantum computers seemed useful for one of two things. First, to provide a general-purpose mechanism to simulate a variety of the real quantum systems that people care about, such as various interactions inside molecules in quantum chemistry. Second, as a challenge to the Physical Extended Church Turing Thesis which says that every physically realizable computation device can be modeled (up to polynomial overhead) by Turing machines (or equivalently, NAND-TM / NAND-RAM programs).

Quantum chemistry is important (and in particular understanding it can be a bottleneck for designing new materials, drugs, and more), but it is still a rather niche area within the broader context of computing (and even scientific computing) applications. Hence for a while most researchers (to the extent they were aware of it), thought of quantum computers as a theoretical curiosity that has little bearing to practice, given that this theoretical “extra power” of quantum computer seemed to offer little advantage in the majority of the problems people want to solve in areas such as combinatorial optimization, machine learning, data structures, etc..

To some extent this is still true today. As far as we know, quantum computers, if built, will not provide exponential speed ups for 95% of the applications of computing. In particular, as far as we know, quantum computers will not help us solve \( \text{NP} \) complete problems in polynomial or even sub-exponential time, though Grover’s algorithm (Remark 22.4) does yield a quadratic advantage in many cases.

However, there is one cryptography-sized exception: In 1994 Peter Shor showed that quantum computers can solve the integer factoring and discrete logarithm in polynomial time. This result has captured the imagination of a great many people, and completely energized re-

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6 As its title suggests, Feynman’s lecture was actually focused on the other side of simulating physics with a computer. However, he mentioned that as a “side remark” one could wonder if it’s possible to simulate physics with a new kind of computer - a “quantum computer” which would “not [be] a Turing machine, but a machine of a different kind”. As far as I know, Feynman did not suggest that such a computer could be useful for computations completely outside the domain of quantum simulation. Indeed, he was more interested in the question of whether quantum mechanics could be simulated by a classical computer.

7 This “95 percent” is a figure of speech, but not completely so. At the time of this writing, cryptocurrency mining electricity consumption is estimated to use up at least 70Twh or 0.3 percent of the world’s production, which is about 2 to 5 percent of the total energy usage for the computing industry. All the current cryptocurrencies will be broken by quantum computers. Also for many web servers the TLS protocol (which is based on the current non-lattice based systems would be completely broken by quantum computing) is responsible for about 1 percent of the CPU usage.
search into quantum computing. This is both because the hardness of these particular problems provides the foundations for securing such a huge part of our communications (and these days, our economy), as well as it was a powerful demonstration that quantum computers could turn out to be useful for problems that a-priori seemed to have nothing to do with quantum physics.

As we’ll discuss later, at the moment there are several intensive efforts to construct large scale quantum computers. It seems safe to say that, as far as we know, in the next five years or so there will not be a quantum computer large enough to factor, say, a 1024 bit number. On the other hand, it does seem quite likely that in the very near future there will be quantum computers which achieve some task exponentially faster than the best-known way to achieve the same task with a classical computer. When and if a quantum computer is built that is strong enough to break reasonable parameters of Diffie Hellman, RSA and elliptic curve cryptography is anybody’s guess. It could also be a “self destroying prophecy” whereby the existence of a small-scale quantum computer would cause everyone to shift away to lattice-based crypto which in turn will diminish the motivation to invest the huge resources needed to build a large scale quantum computer.8

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8 Of course, given that we’re still hearing of attacks exploiting “export grade” cryptography that was supposed to disappear in 1990’s, I imagine that we’ll still have products running 1024 bit RSA when everyone has a quantum laptop.

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22.6 QUANTUM SYSTEMS

Before we talk about quantum computing, let us recall how we physically realize “vanilla” or classical computing. We model a logical bit that can equal 0 or a 1 by some physical system that can be in one of two states. For example, it might be a wire with high or low voltage,
charged or uncharged capacitor, or even (as we saw) a pipe with or without a flow of water, or the presence or absence of a soldier crab. A classical system of \( n \) bits is composed of \( n \) such “basic systems”, each of which can be in either a “zero” or “one” state. We can model the state of such a system by a string \( s \in \{0, 1\}^n \). If we perform an operation such as writing to the 17-th bit the NAND of the 3rd and 5th bits, this corresponds to applying a local function to \( s \) such as setting \( s_{17} = 1 - s_3 \cdot s_5 \).

In the probabilistic setting, we would model the state of the system by a distribution. For an individual bit, we could model it by a pair of non-negative numbers \( \alpha, \beta \) such that \( \alpha + \beta = 1 \), where \( \alpha \) is the probability that the bit is zero and \( \beta \) is the probability that the bit is one. For example, applying the negation (i.e., NOT) operation to this bit corresponds to mapping the pair \((\alpha, \beta)\) to \((\beta, \alpha)\) since the probability that NOT(\( \sigma \)) is equal to 1 is the same as the probability that \( \sigma \) is equal to 0. This means that we can think of the NOT function as the linear map \( N : \mathbb{R}^2 \rightarrow \mathbb{R}^2 \) such that \( N \begin{pmatrix} \alpha \\ \beta \end{pmatrix} = \begin{pmatrix} \beta \\ \alpha \end{pmatrix} \) or equivalently as the matrix \( \begin{pmatrix} 0 & 1 \\ 1 & 0 \end{pmatrix} \).

If we think of the \( n \)-bit system as a whole, then since the \( n \) bits can take one of \( 2^n \) possible values, we model the state of the system as a vector \( p \) of \( 2^n \) probabilities. For every \( s \in \{0, 1\}^n \), we denote by \( e_s \) the \( 2^n \) dimensional vector that has 1 in the coordinate corresponding to \( s \) (identifying it with a number in \( [2^n] \)), and so can write \( p \) as \( \sum_{s \in \{0, 1\}^n} p_s e_s \) where \( p_s \) is the probability that the system is in the state \( s \).

Applying the operation above of setting the 17-th bit to the NAND of the 3rd and 5th bits, corresponds to transforming the vector \( p \) to the vector \( Fp \) where \( F : \mathbb{R}^{2^n} \rightarrow \mathbb{R}^{2^n} \) is the linear map that maps \( e_s \) to \( e_{s_0 \cdots s_3 s_5 s_{16} \cdots s_{n-1}} \).

Please make sure you understand why performing the operation will take a system in state \( p \) to a system in the state \( Fp \). Understanding the evolution of probabilistic systems is a prerequisite to understanding the evolution of quantum systems.

If your linear algebra is a bit rusty, now would be a good time to review it, and in particular make sure you are comfortable with the notions of matrices, vectors, (orthogonal and orthonormal) bases, and norms.
22.6.1 Quantum amplitudes

In the quantum setting, the state of an individual bit (or "qubit", to use quantum parlance) is modeled by a pair of numbers \( \alpha, \beta \) such that \( |\alpha|^2 + |\beta|^2 = 1 \). While in general these numbers can be complex, for the rest of this chapter, we will often assume they are real (though potentially negative), and hence often drop the absolute value operator. (This turns out not to make much of a difference in explanatory power.) As before, we think of \( \alpha^2 \) as the probability that the bit equals 0 and \( \beta^2 \) as the probability that the bit equals 1. As we did before, we can model the NOT operation by the map \( N : \mathbb{R}^2 \to \mathbb{R}^2 \) where \( N(\alpha, \beta) = (\beta, \alpha) \).

Following quantum tradition, instead of using \( e_0 \) and \( e_1 \) as we did above, from now on we will denote the vector \( (1, 0) \) by \( |0\rangle \) and the vector \( (0, 1) \) by \( |1\rangle \) (and moreover, think of these as column vectors). This is known as the Dirac "ket" notation. This means that NOT is the unique linear map \( N : \mathbb{R}^2 \to \mathbb{R}^2 \) that satisfies \( N|0\rangle = |1\rangle \) and \( N|1\rangle = |0\rangle \). In other words, in the quantum case, as in the probabilistic case, NOT corresponds to the matrix

\[
N = \begin{pmatrix} 0 & 1 \\ 1 & 0 \end{pmatrix}.
\] (22.3)

In classical computation, we typically think that there are only two operations that we can do on a single bit: keep it the same or negate it. In the quantum setting, a single bit operation corresponds to any linear map \( OP : \mathbb{R}^2 \to \mathbb{R}^2 \) that is norm preserving in the sense that for every \( \alpha, \beta \), if we apply \( OP \) to the vector \( \begin{pmatrix} \alpha \\ \beta \end{pmatrix} \) then we obtain a vector \( \begin{pmatrix} \alpha' \\ \beta' \end{pmatrix} \) such that \( \alpha'^2 + \beta'^2 = \alpha^2 + \beta^2 \). Such a linear map \( OP \) corresponds to a unitary two by two matrix.\(^{10}\) Keeping the bit the same corresponds to the matrix \( I = \begin{pmatrix} 1 & 0 \\ 0 & 1 \end{pmatrix} \) and (as we’ve seen) the NOT operations corresponds to the matrix \( N = \begin{pmatrix} 0 & 1 \\ 1 & 0 \end{pmatrix} \). But there are other operations we can use as well. One such useful operation is the Hadamard operation, which corresponds to the matrix

\[
H = \frac{1}{\sqrt{2}} \begin{pmatrix} +1 & +1 \\ +1 & -1 \end{pmatrix}.
\] (22.4)

In fact it turns out that Hadamard is all that we need to add to a classical universal basis to achieve the full power of quantum computing.

\(^{10}\) As we mentioned, quantum mechanics actually models states as vectors with complex coordinates. However, this does not make any qualitative difference to our discussion.
22.6.2 Quantum systems: an executive summary

If you ignore the physics and philosophy, for the purposes of understanding the model of quantum computers, all you need to know about quantum systems is the following. The state of a quantum system of n qubits is modeled by an $2^n$ dimensional vector $\psi$ of unit norm (i.e., squares of all coordinates sums up to 1), which we write as $\psi = \sum_{x \in \{0,1\}^n} \psi_x |x\rangle$ where $|x\rangle$ is the column vector that has 0 in all coordinates except the one corresponding to $x$ (identifying $\{0,1\}^n$ with the numbers $\{0, \ldots , 2^n - 1\}$). We use the convention that if $a$, $b$ are strings of lengths $k$ and $\ell$ respectively then we can write the $2^{k+\ell}$ dimensional vector with 1 in the $ab$-th coordinate and zero elsewhere not just as $|ab\rangle$ but also as $|a\rangle|b\rangle$. In particular, for every $x \in \{0,1\}^n$, we can write the vector $|x\rangle$ also as $|x_0\rangle |x_1\rangle \cdots |x_{n-1}\rangle$. This notation satisfies certain nice distributive laws such as $|a\rangle (|b\rangle + |b'\rangle) |c\rangle = |abc\rangle + |ab'c\rangle$.

A quantum operation on such a system is modeled by a $2^n \times 2^n$ unitary matrix $U$ (one that satisfies $UU^\dagger = I$ where $U^\dagger$ is the transpose operation, or conjugate transpose for complex matrices). If the system is in state $\psi$ and we apply to it the operation $U$, then the new state of the system is $U\psi$.

When we measure an $n$-qubit system in a state $\psi = \sum_{x \in \{0,1\}^n} \psi_x |x\rangle$, then we observe the value $x \in \{0,1\}^n$ with probability $|\psi_x|^2$. In this case, the system collapses to the state $|x\rangle$.

22.7 ANALYSIS OF BELL’S INEQUALITY (OPTIONAL)

Now that we have the notation in place, we can show a strategy for Alice and Bob to display “quantum telepathy” in Bell’s Game. Recall that in the classical case, Alice and Bob can succeed in the “Bell Game” with probability at most $3/4 = 0.75$. We now show that quantum mechanics allows them to succeed with probability at least $0.8$.\footnote{The strategy we show is not the best one. Alice and Bob can in fact succeed with probability $\cos^2(\pi/8) \sim 0.854$.}

Lemma 22.5 There is a 2-qubit quantum state $\psi \in \mathbb{C}^4$ so that if Alice has access to the first qubit of $\psi$, can manipulate and measure it and output $a \in \{0,1\}$ and Bob has access to the second qubit of $\psi$ and can manipulate and measure it and output $b \in \{0,1\}$ then $\Pr[a \oplus b = x \land y] \geq 0.8$.

Proof. Alice and Bob will start by preparing a 2-qubit quantum system in the state

$$\psi = \frac{1}{\sqrt{2}} |00\rangle + \frac{1}{\sqrt{2}} |11\rangle \quad (22.5)$$

(this state is known as an EPR pair). Alice takes the first qubit of the system to her room, and Bob takes the second qubit to his room. Now, when Alice receives $x$ if $x = 0$ she does nothing and if $x = 1$ she applies the unitary map $R_{-\pi/8}$ to her qubit where...
\[ R_\theta = \begin{pmatrix} \cos \theta & -\sin \theta \\ \sin \theta & \cos \theta \end{pmatrix} \] is the unitary operation corresponding to rotation in the plane with angle \( \theta \). When Bob receives \( y \), if \( y = 0 \) he does nothing and if \( y = 1 \) he applies the unitary map \( R_{\pi/8} \) to his qubit. Then each one of them measures their qubit and sends this as their response.

Recall that to win the game Bob and Alice want their outputs to be more likely to differ if \( x = y = 1 \) and to be more likely to agree otherwise. We will split the analysis in one case for each of the four possible values of \( x \) and \( y \).

**Case 1:** \( x = 0 \) and \( y = 0 \). If \( x = y = 0 \) then the state does not change. Because the state \( \psi \) is proportional to \( |00\rangle + |11\rangle \), the measurements of Bob and Alice will always agree (if Alice measures 0 then the state collapses to \( |00\rangle \) and so Bob measures 0 as well, and similarly for 1). Hence in the case \( x = y = 1 \), Alice and Bob always win.

**Case 2:** \( x = 0 \) and \( y = 1 \). If \( x = 0 \) and \( y = 1 \) then after Alice measures her bit, if she gets 0 then the system collapses to the state \( |00\rangle \), in which case after Bob performs his rotation, his qubit is in the state \( \cos(\pi/8)|0\rangle + \sin(\pi/8)|1\rangle \). Thus, when Bob measures his qubit, he will get 0 (and hence agree with Alice) with probability \( \cos^2(\pi/8) \geq 0.85 \). Similarly, if Alice gets 1 then the system collapses to \( |11\rangle \), in which case after rotation Bob’s qubit will be in the state \( -\sin(\pi/8)|0\rangle + \cos(\pi/8)|1\rangle \) and so once again he will agree with Alice with probability \( \cos^2(\pi/8) \).

The analysis for **Case 3**, where \( x = 1 \) and \( y = 0 \), is completely analogous to Case 2. Hence Alice and Bob will agree with probability \( \cos^2(\pi/8) \) in this case as well.\(^{12}\)

**Case 4:** \( x = 1 \) and \( y = 1 \). For the case that \( x = 1 \) and \( y = 1 \), after both Alice and Bob perform their rotations, the state will be proportional to

\[ R_{-\pi/8}|0\rangle R_{\pi/8}|0\rangle + R_{-\pi/8}|1\rangle R_{\pi/8}|1\rangle. \quad (22.6) \]

Intuitively, since we rotate one state by 45 degrees and the other state by -45 degrees, they will become orthogonal to each other, and the measurements will behave like independent coin tosses that agree with probability 1/2. However, for the sake of completeness, we now show the full calculation.

\(^{12}\) We are using the (not too hard) observation that the result of this experiment is the same regardless of the order in which Alice and Bob apply their rotations and measurements.
Opening up the coefficients and using \( \cos(-x) = \cos(x) \) and \( \sin(-x) = -\sin(x) \), we can see that (22.6) is proportional to

\[
\begin{align*}
\cos^2(\pi/8)|00\rangle + \cos(\pi/8)\sin(\pi/8)|01\rangle \\
- \sin(\pi/8)\cos(\pi/8)|10\rangle + \sin^2(\pi/8)|11\rangle \\
- \sin^2(\pi/8)|00\rangle + \sin(\pi/8)\cos(\pi/8)|01\rangle \\
- \cos(\pi/8)\sin(\pi/8)|10\rangle + \cos^2(\pi/8)|11\rangle.
\end{align*}
\]

(22.7)

Using the trigonometric identities \( 2\sin(\alpha)\cos(\alpha) = \sin(2\alpha) \) and \( \cos^2(\alpha) - \sin^2(\alpha) = \cos(2\alpha) \), we see that the probability of getting any one of \(|00\rangle, |10\rangle, |01\rangle, |11\rangle \) is proportional to \( \cos(\pi/4) = \sin(\pi/4) = 1/\sqrt{2} \). Hence all four options for \((a, b)\) are equally likely, which mean that in this case \( a = b \) with probability 0.5.

Taking all the four cases together, the overall probability of winning the game is \( \frac{1}{4} \cdot 1 + \frac{1}{2} \cdot 0.85 + \frac{1}{4} \cdot 0.5 = 0.8 \).

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**Remark 22.6 — Quantum vs probabilistic strategies.** It is instructive to understand what is it about quantum mechanics that enabled this gain in Bell’s Inequality. For this, consider the following analogous probabilistic strategy for Alice and Bob. They agree that each one of them output 0 if he or she get 0 as input and outputs 1 with probability \( p \) if they get 1 as input. In this case one can see that their success probability would be \( \frac{1}{2} \cdot 1 + \frac{1}{2}(1 - p) + \frac{1}{2}[2p(1 - p)] = 0.75 - 0.5p^2 \leq 0.75 \).

The quantum strategy we described above can be thought of as a variant of the probabilistic strategy for parameter \( p \) set to \( \sin^2(\pi/8) = 0.15 \). But in the case \( x = y = 1 \), instead of disagreeing only with probability \( 2p(1 - p) = 1/4 \), the existence of the so called “negative probabilities” in the quantum world allowed us to rotate the state in opposing directions to achieve destructive interference and hence a higher probability of disagreement, namely \( \sin^2(\pi/4) = 0.5 \).

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**22.8 QUANTUM COMPUTATION**

Recall that in the classical setting, we modeled computation as obtained by a sequence of basic operations. We had two types of computational models:

- **Non uniform models of computation** such as Boolean circuits and NAND-CIRC programs, where a finite function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) is computable in size \( T \) if it can be expressed as a combination of
$T$ basic operations (gates in a circuit or lines in a NAND-CIRC program)

- *Uniform models of computation* such as Turing machines and NAND-TM programs, where an infinite function $F : \{0, 1\}^* \to \{0, 1\}$ is computable in time $T(n)$ if there is a single algorithm that on input $x \in \{0, 1\}^n$ evaluates $F(x)$ using at most $T(n)$ basic steps.

When considering *efficient computation*, we defined the class $\mathbb{P}$ to consist of all infinite functions $F : \{0, 1\}^* \to \{0, 1\}$ that can be computed by a Turing machine or NAND-TM program in time $p(n)$ for some polynomial $p(\cdot)$. We defined the class $\mathbb{P}_{\text{poly}}$ to consists of all infinite functions $F : \{0, 1\}^* \to \{0, 1\}$ such that for every $n$, the restriction $F|_n$ of $F$ to $\{0, 1\}^n$ can be computed by a Boolean circuit or NAND-CIRC program of size at most $p(n)$ for some polynomial $p(\cdot)$.

We will do the same for *quantum computation*, focusing mostly on the *non uniform* setting of quantum circuits, since that is simpler, and already illustrates the important differences with classical computing.

### 22.8.1 Quantum circuits

A *quantum circuit* is analogous to a Boolean circuit, and can be described as a directed acyclic graph. One crucial difference is that the *out degree* of every vertex in a quantum circuit is at most one. This is because we cannot “reuse” quantum states without measuring them (which collapses their “probabilities”). Therefore, we cannot use the same qubit as input for two different gates.\(^\text{13}\) Another more technical difference is that to express our operations as unitary matrices, we will need to make sure all our gates are *reversible*. This is not hard to ensure. For example, in the quantum context, instead of thinking of NAND as a (non reversible) map from $\{0, 1\}^2$ to $\{0, 1\}$, we will think of it as the reversible map on three qubits that maps $a, b, c$ to $a, b, c \oplus \text{NAND}(a, b)$ (i.e., flip the last bit if NAND of the first two bits is 1). Equivalently, the NAND operation corresponds to the $8 \times 8$ unitary matrix $U_{\text{NAND}}$ such that (identifying $\{0, 1\}^3$ with $[8]$) for every $a, b, c \in \{0, 1\}$, if $|abc\rangle$ is the basis element with 1 in the $abc$-th coordinate and zero elsewhere, then $U_{\text{NAND}}|abc\rangle = |ab(c \oplus \text{NAND}(a, b))\rangle$.\(^\text{14}\) If we order the rows and

\(^\text{13}\) This is known as the *No Cloning Theorem*.

\(^\text{14}\) Readers familiar with quantum computing should note that $U_{\text{NAND}}$ is a close variant of the so called *Toffoli gate* and so QNAND-CIRC programs correspond to quantum circuits with the Hadamard and Toffoli gates.
columns as 000, 001, 010, … , 111, then $U_{NAND}$ can be written as the following matrix:

$$U_{NAND} = \begin{pmatrix}
0 & 1 & 0 & 0 & 0 & 0 & 0 & 0 \\
1 & 0 & 0 & 0 & 0 & 0 & 0 & 0 \\
0 & 0 & 0 & 1 & 0 & 0 & 0 & 0 \\
0 & 0 & 1 & 0 & 0 & 0 & 0 & 0 \\
0 & 0 & 0 & 0 & 1 & 0 & 0 & 0 \\
0 & 0 & 0 & 0 & 0 & 1 & 0 & 0 \\
0 & 0 & 0 & 0 & 0 & 0 & 1 & 0 \\
0 & 0 & 0 & 0 & 0 & 0 & 0 & 1
\end{pmatrix} \quad (22.8)$$

If we have an $n$ qubit system, then for $i, j, k \in [n]$, we will denote by $U_{NAND}^{i,j,k}$ as the $2^n \times 2^n$ unitary matrix that corresponds to applying $U_{NAND}$ to the $i$-th, $j$-th, and $k$-th bits, leaving the others intact. That is, for every $v = \sum_{x \in \{0,1\}^n} v_x |x\rangle$, $U_{NAND}^{i,j,k}v = \sum_{x \in \{0,1\}^n} v_x |x_0 \cdots x_{k-1} (x_k \oplus NAND(x_i, x_j)) x_{k+1} \cdots x_{n-1}\rangle$.

As mentioned above, we will also use the Hadamard or $HAD$ operation. A quantum circuit is obtained by applying a sequence of $U_{NAND}$ and $HAD$ gates, where a $HAD$ gates corresponding to applying the matrix

$$H = \frac{1}{\sqrt{2}} \begin{pmatrix} +1 & +1 \\ +1 & -1 \end{pmatrix} \quad (22.9)$$

Another way to write define $H$ is that for $b \in \{0,1\}$, $H|b\rangle = \frac{1}{\sqrt{2}}|0\rangle + \frac{1}{\sqrt{2}}(-1)^b|1\rangle$. We define $HAD^i$ to be the $2^n \times 2^n$ unitary matrix that applies $HAD$ to the $i$-th qubit and leaves the others intact. Using the ket notation, we can write this as

$$HAD^i \sum_{x \in \{0,1\}^n} v_x |x\rangle = \frac{1}{\sqrt{2}} \sum_{x \in \{0,1\}^n} |x_0 \cdots x_{i-1}\rangle (|0\rangle + (-1)^{x_i}|1\rangle) |x_i \cdots x_{n-1}\rangle \quad (22.10)$$

A quantum circuit is obtained by composing these basic operations on some $m$ qubits. If $m \geq n$, we use a circuit to compute a function $f : \{0,1\}^n \rightarrow \{0,1\}$:

- On input $x$, we initialize the system to hold $x_0, \ldots, x_{n-1}$ in the first $n$ qubits, and initialize all remaining $m-n$ qubits to zero.
- We execute each elementary operation one by one: at every step we apply to the current state either an operation of the form $U_{NAND}^{i,j,k}$ or an operation of the form $HAD^i$ for $i, j, k \in [m]$.
- At the end of the computation, we measure the system, and output the result of the last qubit (i.e. the qubit in location $m-1$).\(^{15}\)
- We say that the circuit computes the function $f$ if the probability that this output equals $f(x)$ is at least $2/3$. Note that this probability

\(^{15}\) For simplicity we restrict attention to functions with a single bit of output, though the definition of quantum circuits naturally extends to circuits with multiple outputs.
is obtained by summing up the squares of the amplitudes of all coordinates in the final state of the system corresponding to vectors $|y\rangle$ where $y_{m-1} = f(x)$.

Formally we define quantum circuits as follows:

**Definition 22.7 — Quantum circuit.** Let $s \geq m \geq n$. A quantum circuit of $n$ inputs, $m - n$ auxiliary bits, and $s$ gates over the $\{U_{NAND}, HAD\}$ basis is a sequence of $s$ unitary $2^m \times 2^m$ matrices $U_0, \ldots, U_{s-1}$ such that each matrix $U_i$ is either of the form $NAND_{i,j,k}$ for $i,j,k \in [m]$ or $HAD^i$ for $i \in [m]$.

A quantum circuit computes a function $f : \{0,1\}^n \rightarrow \{0,1\}$ if the following is true for every $x \in \{0,1\}^n$:

Let $v$ be the vector $v = U_{s-1}U_{s-2}\cdots U_1U_0|x(0^{m-n})\rangle$ (22.11) and write $v$ as $\sum_{y \in \{0,1\}^m} v_y |y\rangle$. Then

$$\sum_{y \in \{0,1\}^m \text{ s.t. } y_{m-1} = f(x)} |v_y|^2 \geq \frac{2}{3}.$$  (22.12)

**P** Please stop here and see that this definition makes sense to you.

Once we have the notion of quantum circuits, we can define the quantum analog of $P_{poly}$ (i.e., the class of functions computable by polynomial size quantum circuits) as follows:

**Definition 22.8 — $BQP_{poly}$.** Let $F : \{0,1\}^* \rightarrow \{0,1\}$. We say that $F \in BQP_{poly}$ if there exists some polynomial $p : \mathbb{N} \rightarrow \mathbb{N}$ such that for every $n \in \mathbb{N}$, if $F_{|n}$ is the restriction of $F$ to inputs of length $n$, then there is a quantum circuit of size at most $p(n)$ that computes $F_{|n}$.

**R** Remark 22.9 — The obviously exponential fallacy. A priori it might seem “obvious” that quantum computing is exponentially powerful, since to perform a quantum computation on $n$ bits we need to maintain the $2^n$ dimensional state vector and apply $2^n \times 2^n$ matrices to it. Indeed popular descriptions of quantum computing (too) often say something along the lines...
that the difference between quantum and classical computer is that a classical bit can either be zero or one while a qubit can be in both states at once, and so in many qubits a quantum computer can perform exponentially many computations at once.

Depending on how you interpret it, this description is either false or would apply equally well to probabilistic computation, even though we’ve already seen that every randomized algorithm can be simulated by a similar-sized circuit, and in fact we conjecture that \( \text{BPP} = \text{P} \).

Moreover, this “obvious” approach for simulating a quantum computation will take not just exponential time but \textit{exponential space} as well, while can be shown that using a simple recursive formula one can calculate the final quantum state using \textit{polynomial space} (in physics this is known as “Feynman path integrals”). So, the exponentially long vector description by itself does not imply that quantum computers are exponentially powerful. Indeed, we cannot prove that they are (i.e., we have not been able to rule out the possibility that every QNAND-CIRC program could be simulated by a NAND-CIRC program/Boolean circuit with polynomial overhead), but we do have some problems (integer factoring most prominently) for which they do provide exponential speedup over the currently best known classical (deterministic or probabilistic) algorithms.

\[ \] 22.8.2 QNAND-CIRC programs (optional)

Just like in the classical case, there is an equivalence between circuits and straight-line programs, and so we can define the programming language QNAND that is the quantum analog of our NAND-CIRC programming language. To do so, we only add a single operation: \( \text{HAD}(\text{foo}) \) which applies the single-bit operation \( H \) to the variable \( \text{foo} \).

We also use the following interpretation to make NAND reversible: \( \text{foo} = \text{NAND}(\text{bar}, \text{blah}) \) means that we modify \( \text{foo} \) to be the XOR of its original value and the NAND of \( \text{bar} \) and \( \text{blah} \). (In other words, apply the 8 by 8 unitary transformation \( U_{\text{NAND}} \) defined above to the three qubits corresponding to \( \text{foo} \), \( \text{bar} \) and \( \text{blah} \).) If \( \text{foo} \) is initialized to zero then this makes no difference.

If \( P \) is a QNAND-CIRC program with \( n \) input variables, \( \ell \) workspace variables, and \( m \) output variables, then running it on the input \( x \in \{0, 1\}^n \) corresponds to setting up a system with \( n + m + \ell \) qubits and performing the following process:

1. We initialize the input variables \( x[0] \ldots x[n-1] \) to \( x_0, \ldots, x_{n-1} \) and all other variables to 0.
2. We execute the program line by line, applying the corresponding physical operation $H$ or $U_{\text{NAND}}$ to the qubits that are referred to by the line.

3. We measure the output variables $Y[0], ..., Y[m-1]$ and output the result (if there is more than one output then we measure more variables).

### 22.8.3 Uniform computation

Just as in the classical case, we can define uniform computational models. For example, we can define the QNAND-TM programming language to be QNAND augmented with loops and arrays just like NAND-TM is obtained from NAND. Using this we can define the class $\text{BQP}$ which is the uniform analog of $\text{BQP}_{/\text{poly}}$. Just as in the classical setting it holds that $\text{BPP} \subseteq P_{/\text{poly}}$ in the quantum setting it can be shown that $\text{BQP} \subseteq \text{BQP}_{/\text{poly}}$. Just like the classical case, we can also use Quantum Turing Machines instead of QNAND-TM to define $\text{BQP}$.

Yet another way to define $\text{BQP}$ is the following: a function $F : \{0, 1\}^* \rightarrow \{0, 1\}$ is in $\text{BQP}$ if (1) $F \in \text{BQP}_{/\text{poly}}$ and (2) moreover for every $n$, the quantum circuit that verifies this can be generated by a classical polynomial time NAND-TM program (or, equivalently, a polynomial-time Turing machine).\[16\] We use this definition here, though an equivalent one can be made using QNAND-TM or quantum Turing machines:

**Definition 22.10 — The class $\text{BQP}$**. Let $F : \{0, 1\}^* \rightarrow \{0, 1\}$. We say that $F \in \text{BQP}$ if there exists a polynomial time NAND-TM program $P$ such that for every $n$, $P(1^n)$ is the description of a quantum circuit $C_n$ that computes the restriction of $F$ to $\{0, 1\}^n$.

The relation between $\text{NP}$ and $\text{BQP}$ is not known (see also Remark 22.4). It is widely believed that $\text{NP} \nsubseteq \text{BQP}$, but there is no consensus whether or not $\text{BQP} \subseteq \text{NP}$. It is quite possible that these two classes are incomparable, in the sense that $\text{NP} \nsubseteq \text{BQP}$ (and in particular no $\text{NP}$-complete function belongs to $\text{BQP}$) but also $\text{BQP} \nsubseteq \text{NP}$ (and there are some interesting candidates for such problems).
It can be shown that \texttt{QNANDEVAL} (evaluating a quantum circuit on an input) is computable by a polynomial size \texttt{QNAND-CIRC} program, and moreover this program can even be generated \textit{uniformly} and hence \texttt{QNANDEVAL} is in \textbf{BQP}. This allows us to “port” many of the results of classical computational complexity into the quantum realm as well.

\begin{remark}
\textbf{Remark 22.11 — Restricting attention to circuits.} Because the non uniform model is a little cleaner to work with, in the rest of this chapter we mostly restrict attention to this model, though all the algorithms we discuss can be implemented in uniform computation as well.
\end{remark}

\section*{22.9 Physically Realizing Quantum Computation}

To realize quantum computation one needs to create a system with \(n\) independent binary states (i.e., “qubits”), and be able to manipulate small subsets of two or three of these qubits to change their state. While by the way we defined operations above it might seem that one needs to be able to perform arbitrary unitary operations on these two or three qubits, it turns out that there are several choices for \textit{universal sets} - a small constant number of gates that generate all others. The biggest challenge is how to keep the system from being measured and collapsing to a single classical combination of states. This is sometimes known as the \textit{coherence time} of the system. The \textit{threshold theorem} says that there is some absolute constant level of errors \(\tau\) so that if errors are created at every gate at rate smaller than \(\tau\) then we can recover from those and perform arbitrary long computations. (Of course there are different ways to model the errors and so there are actually several threshold \textit{theorems} corresponding to various noise models).

There have been several proposals to build quantum computers:

- \textbf{Superconducting quantum computers} use superconducting electric circuits to do quantum computation. This is the direction where there has been most recent progress towards “beating” classical computers.

- \textbf{Trapped ion quantum computers} Use the states of an ion to simulate a qubit. People have made some recent advances on these computers too. While it’s not at all clear that’s the right measuring yard, the current best implementation of Shor’s algorithm (for factoring 15) is done using an ion-trap computer.

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{superconducting_quantum_computer_prototype.png}
\caption{Superconducting quantum computer prototype at Google. Image credit: Google / MIT Technology Review.}
\end{figure}
• **Topological quantum computers** use a different technology. Topological qubits are more stable by design and hence error correction is less of an issue, but constructing them is extremely challenging.

These approaches are not mutually exclusive and it could be that ultimately quantum computers are built by combining all of them together. In the near future, it seems that we will not be able to achieve full fledged large scale universal quantum computers, but rather more restricted machines, sometimes called “Noisy Intermediate-Scale Quantum Computers” or “NISQ”. See this article by John Preskill for some of the progress and applications of such more restricted devices.

### 22.10 SHOR’S ALGORITHM: HEARING THE SHAPE OF PRIME FACTORS

Bell’s Inequality is a powerful demonstration that there is something very strange going on with quantum mechanics. But could this “strangeness” be of any use to solve computational problems not directly related to quantum systems? A priori, one could guess the answer is no. In 1994 Peter Shor showed that one would be wrong:

**Theorem 22.12 — Shor’s Algorithm.** There is a polynomial-time quantum algorithm that on input an integer $M$ (represented in base two), outputs the prime factorization of $M$.

Another way to state Theorem 22.12 is that if we define $\text{FACTORING} : \{0, 1\}^* \rightarrow \{0, 1\}$ to be the function that on input a pair of numbers $(M, X)$ outputs 1 if and only if $M$ has a factor $P$ such that $2 \leq P \leq X$, then $\text{FACTORING}$ is in $\text{BQP}$. This is an exponential improvement over the best known classical algorithms, which take roughly $2^{\tilde{O}(n^{1/3})}$ time, where the $\tilde{O}$ notation hides factors that are polylogarithmic in $n$. While we will not prove Theorem 22.12 in this chapter, we will sketch some of the ideas behind the proof.

#### 22.10.1 Period finding

At the heart of Shor’s Theorem is an efficient quantum algorithm for finding **periods** of a given function. For example, a function $f : \mathbb{R} \rightarrow \mathbb{R}$ is **periodic** if there is some $h > 0$ such that $f(x + h) = f(x)$ for every $x$ (e.g., see Fig. 22.4).

Musical notes yield one type of periodic function. When you pull on a string on a musical instrument, it vibrates in a repeating pattern. Hence, if we plot the speed of the string (and so also the speed of the air around it) as a function of time, it will correspond to some periodic function. The length of the period is known as the **wave length** of the note. The **frequency** is the number of times the function repeats itself within a unit of time. For example, the “Middle C” note has

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**Figure 22.4:** Top: A periodic function. Bottom: An a-periodic function.
a frequency of 261.63 Hertz, which means its period is $1/(261.63)$
seconds.

If we play a chord by playing several notes at once, we get a more
complex periodic function obtained by combining the functions of
the individual notes (see Fig. 22.5). The human ear contains many
small hairs, each of which is sensitive to a narrow band of frequencies.
Hence when we hear the sound corresponding to a chord, the hairs in
our ears actually separate it out to the components corresponding to
each frequency.

It turns out that (essentially) every periodic function $f : \mathbb{R} \to \mathbb{R}$
can be decomposed into a sum of simple wave functions (namely
functions of the form $x \mapsto \sin(\theta x)$ or $x \mapsto \cos(\theta x)$). This is known as
the Fourier Transform (see Fig. 22.6). The Fourier transform makes it
easy to compute the period of a given function: it will simply be the
least common multiple of the periods of the constituent waves.

### 22.10.2 Shor’s Algorithm: A bird’s eye view

On input an integer $M$, Shor’s algorithm outputs the prime factoriza-
tion of $M$ in time that is polynomial in $\log M$. The main steps in the
algorithm are the following:

**Step 1: Reduce to period finding.** The first step in the algorithm is to
pick a random $A \in \{0, 1, \ldots, M - 1\}$ and define the function $F_A :$
$\{0, 1\}^m \to \{0, 1\}^m$ as $F_A(x) = A^x \pmod{M}$ where we identify the
string $x \in \{0, 1\}^m$ with an integer using the binary representation,
and similarly represent the integer $A^x \pmod{M}$ as a string. (We will
choose $m$ to be some polynomial in $\log M$ and so in particular $\{0, 1\}^m$
is a large enough set to represent all the numbers in $\{0, 1, \ldots, M - 1\}$).

Some not-too-hard (though somewhat technical) calculations show
that: (1) The function $F_A$ is periodic (i.e., there is some integer $p_A$
such that $F_A(x + p_A) = F_A(x)$ for almost every $x$) and more importantly
(2) If we can recover the period $p_A$ of $F_A$ for several randomly chosen
A’s, then we can recover the factorization of $M$. Hence, factoring $M$
reduces to finding out the period of the function $F_A$. Exercise 22.2
asks you to work out this for the related task of computing the discrete
logarithm (which underlies the security of the Diffie-Hellman key
exchange and elliptic curve cryptography).

**Step 2: Period finding via the Quantum Fourier Transform.** Using a simple
trick known as “repeated squaring”, it is possible to compute the
map $x \mapsto F_A(x)$ in time polynomial in $m$, which means we can also
compute this map using a polynomial number of NAND gates, and so
in particular we can generate in polynomial quantum time a quantum state $\rho$ that is (up to normalization) equal to

$$\sum_{x \in \{0,1\}^m} |x\rangle|F_A(x)\rangle .$$  \hspace{1cm} (22.13)

In particular, if we were to measure the state $\rho$, we would get a random pair of the form $(x, y)$ where $y = F_A(x)$. So far, this is not at all impressive. After all, we did not need the power of quantum computing to generate such pairs: we could simply generate a random $x$ and then compute $F_A(x)$.

Another way to describe the state $\rho$ is that the coefficient of $|x\rangle|y\rangle$ in $\rho$ is proportional to $f_{A,y}(x)$ where $f_{A,y} : \{0,1\}^m \to \mathbb{R}$ is the function such that

$$f_{A,y}(x) = \begin{cases} 1 & y = A^x \mod M \\ 0 & \text{otherwise} \end{cases} .$$ \hspace{1cm} (22.14)

The magic of Shor’s algorithm comes from a procedure known as the Quantum Fourier Transform. It allows to change the state $\rho$ into the state $\hat{\rho}$ where the coefficient of $|x\rangle|y\rangle$ is now proportional to the $x$-th Fourier coefficient of $f_{A,y}$. In other words, if we measure the state $\hat{\rho}$, we will obtain a pair $(x, y)$ such that the probability of choosing $x$ is proportional to the square of the weight of the frequency $x$ in the representation of the function $f_{A,y}$. Since for every $y$, the function $f_{A,y}$ has the period $p_A$, it can be shown that the frequency $x$ will be (almost\(^{18}\)) a multiple of $p_A$. If we make several such samples $y_0, \ldots, y_k$ and obtain the frequencies $x_1, \ldots, x_k$, then the true period $p_A$ divides all of them, and it can be shown that it is going to be in fact the greatest common divisor (g.c.d.) of all these frequencies: a value which can be computed in polynomial time.

As mentioned above, we can recover the factorization of $M$ from the periods of $F_{A_0}, \ldots, F_{A_t}$ for some randomly chosen $A_0, \ldots, A_t$ in $\{0, \ldots, M-1\}$ and $t$ which is polynomial in $\log M$.

The resulting algorithm can be described in a high (and somewhat inaccurate) level as follows:

**Shor’s Algorithm: (sketch)**

**Input:** Number $M \in \mathbb{N}$.

**Output:** Prime factorization of $M$.

**Operations:**

1. Repeat the following $k = poly(\log M)$ number of times:
   a. Choose $A \in \{0, \ldots, M-1\}$ at random, and let $f_A : \mathbb{Z}_M \to \mathbb{Z}_M$ be the map $x \mapsto A^x \mod M$.

\(^{18}\) The “almost” qualifier again appears because the original function was only “almost” periodic, but it turns out this can be handled by using an “approximate greatest common divisor” algorithm instead of a standard g.c.d. below. The latter can be obtained using a tool known as the continued fraction representation of a number.
b. For \( t = \text{poly}(\log M) \), repeat \( t \) times the following step: Quantum Fourier Transform to create a quantum state \( |\psi\rangle \) over \( \text{poly}(\log(m)) \) qubits, such that if we measure \( |\psi\rangle \) we obtain a pair of strings \((j, y)\) with probability proportional to the square of the coefficient corresponding to the wave function \( x \mapsto \cos(x\pi j/M) \) or \( x \mapsto \sin(x\pi j/M) \) in the Fourier transform of the function \( f_{A,y} : \mathbb{Z}_m \rightarrow \{0, 1\} \) defined as \( f_{A,y}(x) = 1 \) iff \( f_A(x) = y \).

c. If \( j_1, \ldots, j_t \) are the coefficients we obtained in the previous step, then the least common multiple of \( M/j_1, \ldots, M/j_t \) is likely to be the period of the function \( f_A \).

2. If we let \( A_0, \ldots, A_{k-1} \) and \( p_0, \ldots, p_{k-1} \) be the numbers we chose in the previous step and the corresponding periods of the functions \( f_{A_0}, \ldots, f_{A_{k-1}} \) then we can use classical results in number theory to obtain from these a non-trivial prime factor \( Q \) of \( M \) (if such exists). We can now run the algorithm again with the (smaller) input \( M/Q \) to obtain all other factors.

Reducing factoring to order finding is cumbersome, but can be done in polynomial time using a classical computer. The key quantum ingredient in Shor’s algorithm is the quantum Fourier transform.

**Remark 22.13 — Quantum Fourier Transform.** Despite its name, the Quantum Fourier Transform does not actually give a way to compute the Fourier Transform of a function \( f : \{0, 1\}^m \rightarrow \mathbb{R} \). This would be impossible to do in time polynomial in \( m \), as simply writing down the Fourier Transform would require \( 2^m \) coefficients. Rather the Quantum Fourier Transform gives a quantum state where the amplitude corresponding to an element (think: frequency) \( h \) is equal to the corresponding Fourier coefficient. This allows to sample from a distribution where \( h \) is drawn with probability proportional to the square of its Fourier coefficient. This is not the same as computing the Fourier transform, but is good enough for recovering the period.

**22.11 QUANTUM FOURIER TRANSFORM (ADVANCED, OPTIONAL)**

The above description of Shor’s algorithm skipped over the implementation of the main quantum ingredient: the Quantum Fourier Transform.
algorithm. In this section we discuss the ideas behind this algorithm. We will be rather brief and imprecise. Remark 22.3 and Section 22.13 contain references to sources of more information about this topic.

To understand the Quantum Fourier Transform, we need to better understand the Fourier Transform itself. In particular, we will need to understand how it applies not just to functions whose input is a real number but to functions whose domain can be any arbitrary commutative group. Therefore we now take a short detour to (very basic) group theory, and define the notion of periodic functions over groups.

Remark 22.14 — Group theory. While we define the concepts we use, some background in group or number theory might be quite helpful for fully understanding this section.

We will not use anything more than the basic properties of finite Abelian groups. Specifically we use the following notions:

- A finite group \( G \) can be thought of as simply a set of elements and some \textit{binary operation} \( * \) on these elements (i.e., if \( g, h \in G \) then \( g * h \) is an element of \( G \) as well).
- The operation \( * \) satisfies the sort of properties that a product operation does, namely, it is \textit{associative} (i.e., \( (g * h) * f = g * (h * f) \)) and there is some element \( 1 \) such that \( g * 1 = g \) for all \( g \), and for every \( g \in G \) there exists an element \( g^{-1} \) such that \( g * g^{-1} = 1 \).
- A group is called \textit{commutative} (also known as \textit{Abelian}) if \( g * h = h * g \) for all \( g, h \in G \).

The Fourier transform is a deep and vast topic, on which we will barely touch upon here. Over the real numbers, the Fourier transform of a function \( f \) is obtained by expressing \( f \) in the form \( \sum \hat{f}(\alpha)\chi_{\alpha} \) where the \( \chi_{\alpha} \)'s are “wave functions” (e.g. sines and cosines). However, it turns out that the same notion exists for every Abelian group \( G \).

Specifically, for every such group \( G \), if \( f \) is a function mapping \( G \) to \( \mathbb{C} \), then we can write \( f \) as

\[
f = \sum_{g \in G} \hat{f}(g)\chi_{g}, \tag{22.15}
\]

where the \( \chi_{g} \)'s are functions mapping \( G \) to \( \mathbb{C} \) that are analogs of the “wave functions” for the group \( G \) and for every \( g \in G \), \( \hat{f}(g) \) is a complex number known as the \textit{Fourier coefficient of} \( f \) \textit{corresponding to} \( g \).\(^{19}\) The representation (22.15) is known as the \textit{Fourier expansion} or
Fourier transform of $f$, the numbers $(\hat{f}(g))_{g \in \mathbb{G}}$ are known as the Fourier coefficients of $f$ and the functions $(\chi_g)_{g \in \mathbb{G}}$ are known as the Fourier characters. The central property of the Fourier characters is that they are homomorphisms of the group into the complex numbers, in the sense that for every $x, x' \in \mathbb{G}$, $\chi_g(x * x') = \chi_g(x) \chi_g(x')$, where $*$ is the group operation. One corollary of this property is that if $\chi_g(h) = 1$ then $\chi_g$ is $h$ periodic in the sense that $\chi_g(x * h) = \chi_g(x)$ for every $x$. It turns out that if $f$ is periodic with minimal period $h$, then the only Fourier characters that have non zero coefficients in the expression (22.15) are those that are $h$ periodic as well. This can be used to recover the period of $f$ from its Fourier expansion.

22.11.1 Quantum Fourier Transform over the Boolean Cube: Simon’s Algorithm

We now describe the simplest setting of the Quantum Fourier Transform: the group $\{0, 1\}^n$ with the XOR operation, which we’ll denote by $(\{0, 1\}^n, \oplus)$. It can be shown that the Fourier transform over $(\{0, 1\}^n, \oplus)$ corresponds to expressing $f : \{0, 1\}^n \to \mathbb{C}$ as

$$ f = \sum_{y \in \{0, 1\}^n} \hat{f}(y) \chi_y $$

(22.16)

where $\chi_y : \{0, 1\}^n \to \mathbb{C}$ is defined as $\chi_y(x) = (-1)^{\sum_i y_i x_i}$, and

$$ \hat{f}(y) = \frac{1}{\sqrt{2^n}} \sum_{x \in \{0, 1\}^n} f(x) (-1)^{\sum_i y_i x_i}. $$

The Quantum Fourier Transform over $(\{0, 1\}^n, \oplus)$ is actually quite simple:

**Theorem 22.15 — QFT Over the Boolean Cube.** Let $\rho = \sum_{x \in \{0, 1\}^n} f(x) |x\rangle$ be a quantum state where $f : \{0, 1\}^n \to \mathbb{C}$ is some function satisfying $\sum_{x \in \{0, 1\}^n} |f(x)|^2 = 1$. Then we can use $n$ gates to transform $\rho$ to the state

$$ \sum_{y \in \{0, 1\}^n} \hat{f}(y) |y\rangle $$

(22.17)

where $f = \sum_y \hat{f}(y) \chi_y$ and $\chi_y : \{0, 1\}^n \to \mathbb{C}$ is the function $\chi_y(x) = -1^{\sum_i x_i y_i}$.  

**Proof Idea:**  
The idea behind the proof is that the Hadamard operation corresponds to the Fourier transform over the group $\{0, 1\}^n$ (with the XOR operations). To show this, we just need to do the calculations.  

•
Proof of Theorem 22.15. We can express the Hadamard operation $HAD$ as follows:

$$HAD|a⟩ = \frac{1}{\sqrt{2}}(|0⟩ + (−1)^a|1⟩).$$ (22.18)

We are given the state

$$\rho = \sum_{x \in \{0,1\}^n} f(x)|x⟩.$$ (22.19)

Now suppose that we apply the $HAD$ operation to each of the $n$ qubits. We can see that we get the state

$$2^{−n/2} \sum_{x \in \{0,1\}^n} f(x)\prod_{i=0}^{n-1} (|0⟩ + (−1)^{x_i}|1⟩).$$ (22.20)

We can now use the distributive law and open up a term of the form

$$f(x)(|0⟩ + (−1)^{x_0}|1⟩)\cdots(|0⟩ + (−1)^{x_{n−1}}|1⟩)$$ (22.21)

to the following sum over $2^n$ terms:

$$f(x)\sum_{y \in \{0,1\}^n} (−1)^{\sum_i y_ix_i}|y⟩.$$ (22.22)

But by changing the order of summations, we see that the final state is

$$\sum_{y \in \{0,1\}^n} 2^{−n/2}(\sum_{x \in \{0,1\}^n} f(x)(−1)^{\sum_i x_iy_i})|y⟩$$ (22.23)

which exactly corresponds to $\hat{\rho}$.

\[\blacksquare\]

22.11.2 From Fourier to Period finding: Simon’s Algorithm (advanced, optional)

Using Theorem 22.15 it is not hard to get an algorithm that can recover a string $h^* \in \{0,1\}^n$ given a circuit that computes a function $F : \{0,1\}^n \rightarrow \{0,1\}^*$ that is $h^*$ periodic in the sense that $F(x) = F(x’)$ for distinct $x, x’$ if and only if $x’ = x ⊕ h^*$. The key observation is that if we compute the state $\sum_{x \in \{0,1\}^n} |x⟩ |F(x)⟩$, and perform the Quantum Fourier transform on the first $n$ qubits, then we would get a state such that the only basis elements with nonzero coefficients would be of the form $|y⟩$ where

$$\sum y_i h_i^* = 0 (mod 2)$$ (22.24)

So, by measuring the state, we can obtain a sample of a random $y$ satisfying (22.24). But since (22.24) is a linear equation modulo 2
about the unknown \( n \) variables \( h_0, \ldots, h_{n-1} \), if we repeat this procedure to get \( n \) such equations, we will have at least as many equations as variables and (it can be shown that) this will suffice to recover \( h^* \).

This result is known as Simon’s Algorithm, and it preceded and inspired Shor’s algorithm.

22.11.3 From Simon to Shor (advanced, optional)

Theorem 22.15 seemed to really use the special bit-wise structure of the group \( \{0, 1\}^n \), and so one could wonder if it can be extended to other groups. However, it turns out that we can in fact achieve such a generalization.

The key step in Shor’s algorithm is to implement the Fourier transform for the group \( \mathbb{Z}_L \), which is the set of numbers \( \{0, \ldots, L-1\} \) with the operation being addition modulo \( L \). In this case it turns out that the Fourier characters are the functions \( \chi_y(x) = \omega^{xy} \) where \( \omega = e^{2\pi i/L} \) (\( i \) here denotes the complex number \( \sqrt{-1} \)). The \( y \)-th Fourier coefficient of a function \( f : \mathbb{Z}_L \to \mathbb{C} \) is

\[
\hat{f}(y) = \frac{1}{\sqrt{L}} \sum_{x \in \mathbb{Z}_L} f(x)\omega^{xy} .
\]  

(22.25)

The key to implementing the Quantum Fourier Transform for such groups is to use the same recursive equations that enable the classical Fast Fourier Transform (FFT) algorithm. Specifically, consider the case that \( L = 2^\ell \). We can separate the sum over \( x \) in (22.25) to the terms corresponding to even \( x \)'s (of the form \( x = 2z \)) and odd \( x \)'s (of the form \( x = 2z + 1 \)) to obtain

\[
\hat{f}(y) = \frac{1}{\sqrt{L}} \sum_{z \in \mathbb{Z}_{L/2}} f(2z)(\omega^2)^yz + \frac{\omega^y}{\sqrt{L}} \sum_{z \in \mathbb{Z}_{L/2}} f(2z + 1)(\omega^2)^yz
\]  

(22.26)

which reduces computing the Fourier transform of \( f \) over the group \( \mathbb{Z}_{2^\ell} \) to computing the Fourier transform of the functions \( f_{\text{even}} \) and \( f_{\text{odd}} \) (corresponding to the applying \( f \) to only the even and odd \( x \)'s respectively) which have \( 2^{\ell-1} \) inputs that we can identify with the group \( \mathbb{Z}_{2^{\ell-1}} = \mathbb{Z}_{L/2} \).

Specifically, the Fourier characters of the group \( \mathbb{Z}_{L/2} \) are the functions \( \chi_y(x) = e^{2\pi i/(L/2)x} = (\omega^2)^{yx} \) for every \( x, y \in \mathbb{Z}_{L/2} \). Moreover, since \( \omega^L = 1 \), \( (\omega^2)^y = (\omega^2)^y \mod L/2 \) for every \( y \in \mathbb{N} \). Thus (22.26) translates into

\[
\hat{f}(y) = \hat{f}_{\text{even}}(y \mod L/2) + \omega^y \hat{f}_{\text{odd}}(y \mod L/2) .
\]  

(22.27)

This observation is usually used to obtain a fast (e.g. \( O(L \log L) \)) time to compute the Fourier transform in a classical setting, but it can be used to obtain a quantum circuit of \( \text{poly}(\log L) \) gates to transform a state of the form \( \sum_{x \in \mathbb{Z}_L} f(x)|x\rangle \) to a state of the form \( \sum_{y \in \mathbb{Z}_L} \hat{f}(y)|y\rangle \).
The case that $L$ is not an exact power of two causes some complications in both the classical case of the Fast Fourier Transform and the quantum setting of Shor’s algorithm. However, it is possible to handle these. The idea is that we can embed $Z_L$ in the group $\mathbb{Z}_{A \cdot L}$ for any integer $A$, and we can find an integer $A$ such that $A \cdot L$ will be close enough to a power of 2 (i.e., a number of the form $2^m$ for some $m$), so that if we do the Fourier transform over the group $\mathbb{Z}_{2^m}$ then we will not introduce too many errors.

### Lecture Recap

- The state of an $n$-qubit quantum system can be modeled as a $2^n$ dimensional vector.
- An operation on the state corresponds to applying a unitary matrix to this vector.
- Quantum circuits are obtained by composing basic operations such as $HAD$ and $U_{NAND}$.
- We can use quantum circuits to define the classes $\text{BQP}_{\text{poly}}$ and $\text{BQP}$ which are the quantum analogs of $\text{P}_{\text{poly}}$ and $\text{BPP}$ respectively.
- There are some problems for which the best known quantum algorithm is exponentially faster than the best known, but quantum computing is not a panacea. In particular, as far as we know, quantum computers could still require exponential time to solve $\text{NP}$-complete problems such as $\text{SAT}$.

### 22.12 Exercises

**Exercise 22.1 — Quantum and classical complexity class relations.** Prove the following relations between quantum complexity classes and classical ones:

1. $\text{P}_{\text{poly}} \subseteq \text{BQP}_{\text{poly}}$. See footnote for hint.\(^{21}\)
2. $\text{P} \subseteq \text{BQP}$. See footnote for hint.\(^{22}\)
3. $\text{BPP} \subseteq \text{BQP}$. See footnote for hint.\(^{23}\)
4. $\text{BQP} \subseteq \text{EXP}$. See footnote for hint.\(^{24}\)
5. If $\text{SAT} \in \text{BQP}$ then $\text{NP} \subseteq \text{BQP}$. See footnote for hint.\(^{25}\)

**Exercise 22.2 — Discrete logarithm from order finding.** Show a probabilistic polynomial time classical algorithm that given an Abelian finite group $G$ (in the form of an algorithm that computes the group operation), a generator $g$ for the group, and an element $h \in G$, as well as access to a black box that on input $f \in G$ outputs the order of $f$ (the smallest $a$

\(^{21}\) You can use $U_{NAND}$ to simulate NAND gates.

\(^{22}\) Use the alternative characterization of $\text{P}$ as in ??.

\(^{23}\) You can use the $HAD$ gate to simulate a coin toss.

\(^{24}\) In exponential time simulating quantum computation boils down to matrix multiplication.

\(^{25}\) If a reduction can be implemented in $\text{P}$ it can be implemented in $\text{BQP}$ as well.
such that $f^a = 1$, computes the discrete logarithm of $h$ with respect to $g$. That is the algorithm should output a number $x$ such that $g^x = h$. See footnote for hint.26

22.13 BIBLIOGRAPHICAL NOTES

Chapters 9 and 10 in the book Quantum Computing Since Democritus give an informal but highly informative introduction to the topics of this lecture and much more. Shor’s and Simon’s algorithms are also covered in Chapter 10 of my book with Arora on computational complexity.

There are many excellent videos available online covering some of these materials. The Fourier transform is covered in these videos of Dr. Chris Geoscience, Clare Zhang and Vi Hart. More specifically to quantum computing, the videos of Umesh Vazirani on the Quantum Fourier Transform and Kelsey Houston-Edwards on Shor’s Algorithm are very recommended.

Chapter 10 in Avi Wigderson’s book gives a high level overview of quantum computing. Andrew Childs’ lecture notes on quantum algorithms, as well as the lecture notes of Umesh Vazirani, John Preskill, and John Watrous.

Regarding quantum mechanics in general, this video illustrates the double slit experiment, this Scientific American video is a nice exposition of Bell’s Theorem. This talk and panel moderated by Brian Greene discusses some of the philosophical and technical issues around quantum mechanics and its so called “measurement problem”. The Feynmann lecture on the Fourier Transform and quantum mechanics in general are very much worth reading.

The Fast Fourier Transform, used as a component in Shor’s algorithm, is one of the most useful algorithms across many application areas. The stories of its discovery by Gauss in trying to calculate asteroid orbits and rediscovery by Tukey during the cold war are fascinating as well.

22.14 FURTHER EXPLORATIONS

Some topics related to this chapter that might be accessible to advanced students include: (to be completed)

22.15 ACKNOWLEDGEMENTS

Thanks to Scott Aaronson for many helpful comments about this chapter.

26 We are given $h = g^x$ and need to recover $x$. To do so we can compute the order of various elements of the form $h^a g^b$. The order of such an element is a number $c$ satisfying $c(xa + b) = 0$ (mod $|G|$).

With a few random examples we will get a non trivial equation on $x$ (where $c$ is not zero modulo $|G|$) and then we can use our knowledge of $a, b, c$ to recover $x$. 
VI
APPENDICES
Bibliography


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