In early computer systems, much effort was taken to drive out randomness and noise. Hardware components were prone to nondeterministic 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
uncle gambled). The name stuck, and randomized algorithms are known as Monte Carlo algorithms to this day.

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 $\text{NP}$-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 $m/2$.

⋆

**Proof of Theorem 18.1.** The algorithm is extremely simple:

<table>
<thead>
<tr>
<th>Algorithm Random Cut:</th>
</tr>
</thead>
<tbody>
<tr>
<td><strong>Input:</strong> Graph $G = (V, E)$ with $n$ vertices and $m$ edges. Denote $V = {v_0, v_1, \ldots, v_{n-1}}$.</td>
</tr>
<tr>
<td><strong>Operation:</strong></td>
</tr>
<tr>
<td>1. Pick $x$ uniformly at random in ${0, 1}^n$.</td>
</tr>
<tr>
<td>2. 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 $x_i = 1$.</td>
</tr>
<tr>
<td>3. Output the cut $(S, \overline{S})$.</td>
</tr>
</tbody>
</table>

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 \( \mathbb{E}[X_e] = 1/2 \). 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

\[
\mathbb{E}[X] = \sum_{e \in E} \mathbb{E}[X_e] = m(1/2) = m/2 .
\] (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**.

\[ \star \]

**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

\[
\begin{align*}
pm + (1-p)(m/2 - 0.5) &\leq pm + m/2 - 0.5 \\
&= (18.2)
\end{align*}
\]

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 \( m/2 \).

\[\blacksquare\]

**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

\[
(1 - 1/(2m))^{2000m} \leq 2^{-1000}.
\]

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

\[
\Pr[\text{number of edges cut by } (S, \overline{S}) \geq |E|/2] \geq 1 - 2^{-k}.
\]

**(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 \( F : \{0, 1\}^* \rightarrow \{0, 1\} \). 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., \( 1 - 2^{-Ω(k)} \)), 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 \( F : \{0, 1\}^* \rightarrow \{0, 1\} \) is a function such that there is a polynomial-time algorithm \( A \) satisfying

\[
\Pr[A(x) = F(x)] \geq 0.51
\]  

(18.5)

for every \( x \in \{0, 1\}^* \), then there is a polynomial time algorithm \( B \) satisfying

\[
\Pr[B(x) = F(x)] \geq 1 - 2^{-|x|}
\]  

(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
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. 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 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:
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 $O(1.3303^n)$ in the worst case.

a. Choose a random assignment $x \in \{0, 1\}^n$ and repeat the following for $S$ steps:
   1. If $x$ satisfies $\varphi$ then output $x$.
   2. Otherwise, choose a random clause $(\ell_i \lor \ell_j \lor \ell_k)$ that $x$ does not satisfy, choose a random literal in $\ell_i, \ell_j, \ell_k$ and modify $x$ to satisfy this literal.

2. If all the $T \cdot S$ repetitions above did not result in a satisfying assignment then output Unsatisfiable

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 $ST = \tilde{O}((4/3)^n) = \tilde{O}(1.333 \ldots^n)$ (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 II:** 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, $\text{FLIP}(x_0, \ldots, x_{n-1}) = (1 - x_0, \ldots, 1 - x_{n-1})$. 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

---

3 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 $O(1.3303^n)$ in the worst case.
\(|A| = |B|\) 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, they have the same probability and so we get that \(2 \Pr[B] \leq 1\) or \(\Pr[B] \leq 1/2\). (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 \(\pi : [n] \rightarrow [n]\)) 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 \(\mathbb{R}^{n^2} \rightarrow \mathbb{R}\) where

\[
P(x_0,0, ..., x_{n-1,n-1}) = \sum_{\pi \in S_n} \left( \prod_{i=0}^{n-1} \text{sign}(\pi) A_{i,\pi(i)} \right) \prod_{i=0}^{n-1} x_{i,\pi(i)}
\]

(18.7)
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
Theorem 18.7 — Schwartz–Zippel lemma. For every integer \( q \), and polynomial \( P : \mathbb{R}^n \rightarrow \mathbb{R} \) 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.  

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 $\{0, \ldots, M - 1\}$ 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 $F : \{0, 1\}^n \to \{0, \ldots, M - 1\} \cup \{\perp\}$ 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 $\{0, \ldots, M - 1\}$.  

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 $2^{1-H(1/4)+(1/4)\log 3} = (3/2)$.
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).  

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