

Monotone circuit lower bounds from spread matchings

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Abstract

We prove new exponential monotone circuit lower bounds for detecting perfect matchings. We show that $\exp(\tilde{\Omega}(n^{1/2}))$ gates are required to detect bipartite perfect matchings in n -vertex graphs. Our lower bounds are based on a new spread matching lemma.

1. Introduction

A monotone Boolean circuit is a circuit that only uses the monotone gates $\wedge, \vee$; negations are not allowed. One of the beautiful success stories of complexity theory has been the achievement of exponential monotone circuit lower bounds. In this paper, we prove new monotone circuit lower bounds for detecting matchings, and simplify the proofs of some other monotone circuit lower bounds for natural functions.

In a foundational paper, Razborov [20] proved that computing matching and detecting cliques both require monotone circuits of size $n^{\Omega(\log n)}$. Alon and Boppana [1] later improved the lower bound for clique to $\exp(\tilde{\Omega}(n^{1/3}))$. These early results leveraged the combinatorial concept of *sunflowers*, namely a collection of sets whose pairwise intersections are identical. Erdős and Rado [10] showed that every large collection of small sets must contain a large sunflower, and Razborov used this sunflower lemma to approximate the gates of a given monotone circuit with simpler functions, yielding circuit lower bounds. Since then, similar monotone circuit lower bounds were also proved by appealing to the switching lemma or other methods which avoided the use of sunflowers altogether [13, 5, 14]. Another direction leveraged the connection to communication complexity to prove optimal lower bounds of $\Omega(n)$ on the depth of the monotone circuits computing clique or matching [19, 18].

1.1. Robust sunflowers and spreadness

A significant advance in the study of these monotone circuit lower bounds was made by Rossman [21] when he introduced the notion of *robust sunflowers*. Loosely speaking, a robust sunflower is a DNF, namely a formula in disjunctive normal form, with the property

that the DNF can be approximated by a single term. Rossman's work revealed that robust sunflowers could be used to prove new monotone circuit lower bounds. For example, he gave exponential monotone circuit lower bounds for the task of distinguishing graphs with cliques of size k from Erdős-Renyi random graphs that do not have such cliques.

In a breakthrough, Alweiss, Lovett, Wu and Zhang [2] established an essentially optimal robust sunflower lemma. The proof crucially used the concept of a *spread* set, first considered by Talagrand [22]. We say that a random set X is r -spread if for all sets S ,

$$p(S \subseteq X) \leq r^{-|S|}.$$

Let $\mathcal{X}$ denote the support of X . The key contribution of [2] is a nice encoding argument to show that a random set must contain some element of a spread family $\mathcal{X}$ with high probability. Subsequent work by myself [16], Frankston, Kahn, Narayanan and Park [11], and Bell, Chueluecha and Warnke [4] cleaned up the bounds to obtain the following version (see [17] for the easiest proof):

Lemma 1 (Spread Set Lemma). *There is a universal constant $\alpha > 0$ such that the following holds. Let $\varepsilon \leq 1/2$, $r = \alpha \log(t/\varepsilon)/\gamma$ and X be a random r -spread set of size at most t . Then if Y is a random set including each element independently with probability γ ,*

$$p(\exists x \in \mathcal{X}, x \subseteq Y) > 1 - \varepsilon.$$

If a monotone DNF has terms that correspond to $\mathcal{X}$, then the conclusion of Lemma 1 is that the DNF can be approximated by the function 1 from the perspective of a random input Y . Lemma 1 is tight [2]. Let $[n] = \{1, 2, \dots, n\}$, let $f : [t] \rightarrow [\lceil \log(t/\varepsilon)/\gamma \rceil]$ be a uniformly random function, and let $X = \{(i, f(i))\}$ be its graph. Then X is r -spread with $r = \lceil \log(t/\varepsilon)/\gamma \rceil$, but if Y includes each element with probability γ , the probability that Y is disjoint from $[i] \times [\lceil \log(t/\varepsilon)/\gamma \rceil]$ is

$$(1 - \gamma)^{\lceil \log(t/\varepsilon)/\gamma \rceil} \approx e^{\gamma \lceil \log(t/\varepsilon)/\gamma \rceil} \approx \varepsilon/t,$$

so Y does not contain the graph of any f with probability $\approx \varepsilon$.

These results led to several new monotone circuit lower bounds, as well as new connections to so-called *lifting* results by Garg, Göös, Kamath and Sokolov [12], Lovett, Meka, Mertz, Pitassi and Zhang [15] and by de Rezende and Vinyals [9]. For example, Cavalar, Kumar and Rossman [8] used Lemma 1 to prove a lower bound of $\exp(\tilde{\Omega}(\sqrt{n}))$ for an explicit function of n variables, beating the prior record of $\exp(\tilde{\Omega}(n^{1/3}))$ due to Andreev [3] and Harnik and Raz [14]. They also gave a lower bound of $\exp(\tilde{\Omega}(k))$ for detecting cliques up to size $k \leq n^{1/3}$, by proving a *robust clique sunflower lemma*. Blasiok and Meierhöfer [6] and de Rezende and Vinyals [9] proved lower bounds of the type $\exp(\tilde{\Omega}(\sqrt{n}))$ for monotone circuits computing whether or not a graph on n vertices contains a clique of size $\approx \sqrt{n}$. The stronger bounds of [6] followed from a better robust clique sunflower lemma. Implicit in their work is the following *spread clique sunflower lemma* (see Section 4 for the proof). For a set S , let $\mathcal{C}_S$ denote the clique on the vertices of S .

Lemma 2 (Spread Clique Lemma). *There is a universal constant $\alpha > 0$ such that the following holds. Let $\varepsilon \leq 1/2$, $r = \alpha \log(t/\varepsilon)/p^{t+s}$ and $X \subseteq [n]$ be a random r -spread set of size at most t . Then if Y is a graph on the vertex set $[n]$ sampled by including each edge independently with probability p and $S \subseteq [n]$ is an arbitrary set of size s ,*

$$p(\exists x \in \mathcal{X}, \mathcal{C}_{x \cup S} \subseteq Y \cup \mathcal{C}_S) > 1 - \varepsilon.$$

Razborov's $n^{\Omega(\log n)}$ bound remained the best monotone circuit lower bound for matching until Cavalar, Göös, Riazanov, Sofronova, and Sokolov [7] used Lemma 1 as well as the concept of *blocky matchings* to prove a robust matching sunflower lemma and used it to show that any monotone circuit distinguishing graphs with a perfect matching from graphs with maximum matching size $n - 2$ must have size $\exp(\tilde{\Omega}(n^{1/3}))$.

1.2. Our results

In this paper, we give new monotone circuit lower bounds for detecting matchings.

Theorem 3. *For $q < k \leq n/2$, any monotone circuit that can distinguish n -vertex bipartite graphs with matchings of size k from graphs that do not have matchings of size q must have $\exp(\Omega(q/(\sqrt{k} \log n)))$ gates.*

Setting $k = n/2$, $q = n/4$ yields:

Corollary 4. *Any monotone circuit that can distinguish n -vertex bipartite graphs with a perfect matching from graphs that do not have matchings of size $n/4$ must have $\exp(\Omega(\sqrt{n}/\log n))$ gates.*

Corollary 4 improves on the results of [7] in two ways: the exponent is larger, and the lower bound applies for distinguishing perfect matchings from graphs with no small matchings, rather than distinguishing perfect matchings from graphs with matchings of size $n - 2$. Theorem 3 gives exponential lower bounds even for circuits that distinguish graphs with matchings of size $n^{1/5}$ from graphs that do not have matchings of size $n^{1/9}$.

Following the usual pattern, these new results are a consequence of a new *spread matching lemma* with essentially optimal parameters. We say that a random matching X is r -spread if for every matching m ,

$$p(m \subseteq X) \leq r^{-|m|}.$$

Let $G(W)$ denote the graph that contains exactly those edges that are incident to W .

Lemma 5 (Spread Matching Lemma). *There is a universal constant $\alpha > 0$ such that the following holds. Let $\varepsilon < 1/2$, $t > 1$, $r = \alpha \cdot (\log t)^2 \cdot (t + \log^2(1/\varepsilon))/\gamma^2$, and X be a random r -spread matching on the set of vertices $[n]$, with $|X| \leq t$. Then if $W \subseteq [n]$ is a uniformly random set with $|W| = \gamma n$,*

$$p(\exists x \in \mathcal{X}, x \subseteq G(W)) > 1 - \varepsilon.$$

Remark 6. Lemma 5 also holds when W includes each element independently with probability γ . To prove this, add dummy elements to the universe until the probability that W contains $< \gamma n/2$ elements is much less than ε , and then apply Lemma 5.

A consequence of Lemma 5 is the following robust spread matching lemma.

Corollary 7 (Robust Spread Matching Lemma). *There is a universal constant $\alpha > 0$ such that the following holds. Let $\varepsilon < 1/2$, $t > 1$, $r = \alpha \cdot (\log t)^2 \cdot (t + \log^2(1/\varepsilon))/\gamma^2$, and let $\mathcal{X}$ be a family of r^t matchings of size at most t . Then there is a matching m such that if $W \subseteq [n]$ is a uniformly random set with $|W| = \gamma n$,*

$$p(\exists m' \in \mathcal{X}, m \subseteq m', m' \subseteq G(W) \cup m) > 1 - \varepsilon.$$

Corollary 7 looks closer to a statement proved in [7]. To prove it, let $X \in \mathcal{X}$ be uniform, and pick the largest matching m violating r -spreadness. Then $X \setminus m \mid X \in \mathcal{M}$ is guaranteed to be r -spread by the choice of m , so we can apply Lemma 5 to derive the conclusion. The main improvement of Lemma 5 and Corollary 7 over the corresponding statement of [7] is the dependence of r on t and ε . Our dependence is of the type $t + \log^2(1/\varepsilon)$, while theirs is $t \log^2(1/\varepsilon)$. Our results apply using the graph $G(W)$, which contains a matching of size at most γn , while their result required the use of graphs that contain matchings of size $n - 2$.

The proof of Lemma 5 is not as easy as the old one. It uses a new encoding argument tailored for matchings, and the argument is, unfortunately, quite subtle. We prove the lemma in Section 2, where we also give several examples that verify the (near) optimality of its parameters. One of the examples given there shows that $r \geq \Omega(\log^2(1/\varepsilon))$ is necessary to obtain error ε .

In an effort to unify all known monotone circuit lower bounds into a single nice framework, one that we prefer over the other known ways of reasoning about these things, we reprove the results of [8, 6, 9]. For detecting cliques, we use Lemma 2 to prove the following in Section 4:

Theorem 8. *There is a universal constant $\beta > 0$ such that if $q < k$, $\beta k \log q \leq n$, any monotone circuit distinguishing n -vertex graphs that contain a clique of size k from graphs that have no clique of size q must have size $\left(\frac{\beta n}{qk}\right)^{q/\log n}$.*

Setting $k = \beta\sqrt{n}/2$, $q = k/2$ proves a lower bound of $\exp(\Omega(\sqrt{n}/\log n))$, which saves a log-factor over the results of [9]. We reproduce the bound proved by [8] in Section 5. Without loss of generality, assume n is a prime. Let D_d denote the set of univariate polynomials of degree d over $\mathbb{F}_n$. Define

$$F_{k,d}(x) = \bigvee_{\substack{w \in D_d \\ |w([k])| \geq k/2}} \bigwedge_{i=1}^k x_{w(i)}.$$

In Section 5, we prove

Theorem 9. *There is a universal constant $\beta > 0$ such that if $\beta k \log d \leq n$ and $3d \log n \leq k$, then $F_{k,d}$ requires monotone circuits of size $\exp(\Omega(\min\{d, n/k\}))$.*

Setting $k = \sqrt{n}$, $d = k/\log n$ yields $F_{k,d} \in NP$, but $F_{k,d}$ requires monotone circuits of size $\exp(\Omega(\sqrt{n}/\log n))$, matching [8].

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2. Spread matching lemma

In this section, we prove Lemma 1. Here are some examples that show why Lemma 5 has must have the form that it does:

1. Let A and B be disjoint sets, each of size $\log(1/\varepsilon)/2$, and X be a single random edge from A to B . Let $\gamma = 1/2$. Then X is r -spread with $r = \log(1/\varepsilon)^2/4$, yet with probability $\approx \varepsilon$, $G(W)$ does not include any edge in the support of X . This shows that $r \geq \Omega(\log^2(1/\varepsilon))$ is required.
2. Let A and B be disjoint sets of size t , and X be a uniformly random perfect matching from A to B . Then X is r -spread with $r = t/2$. If $\gamma = 1/3$, with probability $> 1/2$, W contains less than t of the vertices of $A \cup B$, so the probability that $G(W)$ contains one of the matchings is at most $1/2$. This shows that $(t/2)$ -spreadness is necessary, even to achieve error $1/2$.
3. Let X be a uniformly random matching of size $t = \lambda n/2$ from $[n/2]$ to $[n] \setminus [n/2]$. Then X is $\Theta(n/\lambda)$ -spread. However, if $\gamma < \lambda/2$, we see that $G(W)$ cannot contain any matching of size t . This shows that $r \ll t/\gamma^2$ is insufficient.
4. Let X be a uniformly random edge from the star graph, namely the set of edges $\{\{1, j\} : j > 1\}$. We see that X is $(n-1)$ -spread, yet W contains an edge of $\mathcal{X}$ only if it contains 1, and this happens with probability at most γ . This example shows why edges not contained in W need to be included in $G(W)$.

2.1. Proof of Lemma 5

Let q be the smallest number such that $2^q \geq t$, and sample a family of uniformly random disjoint sets $\{W_{i,j}\}$, with

$$i \in [q], \quad j \in [\lceil 2^{i-q} \log(q/\varepsilon) \rceil], \quad |W_{i,j}| = \left\lfloor \frac{\gamma n}{q \cdot \lceil 2^{i-q} \log(q/\varepsilon) \rceil} \right\rfloor.$$

Let $W = \cup_{i,j} W_{i,j}$. We have $|W| \leq \gamma n$. Sample $W_{i,j}$ in the order

$$W_{1,1}, W_{1,2}, \dots, W_{1, \lceil 2^{1-q} \log(q/\varepsilon) \rceil}, W_{2,1}, W_{2,2}, \dots, W_{q, \lceil \log(q/\varepsilon) \rceil}.$$

Let $k(i, j)$ be the location of the set $W_{i,j}$ in the above ordering: $k(1, 1) = 1, k(1, 2) = 2$ and so on. We write $W_k = W_{i,j}$ and $W^k = \bigcup_{l=1}^k W_l$. Let $W^0 = W_0 = \emptyset$. During this random process, we need to keep track of the matchings of $\mathcal{X}$ that have been partially covered by $G(W^k)$. To this end, define the map π_k as follows. Set $\pi_0(x) = x$, and

$$\pi_k(x) = \pi_{k-1}(\mu_k(x)) \setminus G(W^k),$$

where $\mu_k(x) \in \mathcal{X}$ is the lexicographically first matching that minimizes $|\pi_{k-1}(\mu_k(x)) \setminus G(W^k)|$ subject to the constraints

$$\pi_{k-1}(\mu_k(x)) \subseteq G(W^k) \cup \pi_{k-1}(x),$$

and

$$|\pi_{k-1}(\mu_k(x))| \leq 2^{q-i+1}.$$

This choice ensures that $\pi_k(x) \subseteq \pi_{k-1}(x)$, and in particular $\pi_k(x) \subseteq x$. The map π_k is determined by $W_1, \dots, W_k$. Given W^k , $\pi_k(x)$ is a matching that is easier to cover than $\pi_{k-1}(x)$, and if $\pi_k(x)$ is eventually covered by $G(W^k)$, then some matching of $\mathcal{X}$ must also be covered. If $\pi_k(x)$ is empty, then $\mathcal{X}$ must contain a matching that has been covered by $G(W^k)$.

For $k = k(i, j)$, let $a \geq 2^{q-i}$, and fix $W_1, \dots, W_{k-1}$. Let E_k denote the event:

$$E_k : \quad |\pi_{k-1}(X)| \leq 2^{q-i+1} \text{ and } |\pi_k(X)| \geq 2^{q-i}.$$

The key technical bound of the proof is that over the random choice of the pair (W_k, X) we have

$$p(E_k) \leq \kappa^{-2^{q-i}}, \tag{1}$$

where here $\kappa = \alpha/2^{17}$, for α as in the statement of Lemma 5.

Let us see how to use Equation (1) to finish the proof. Applying Markov's inequality to Equation (1), we conclude that

$$p_{W_k}(p_X(E_k) > \kappa^{-2^{q-i-1}}) < \kappa^{-2^{q-i-1}},$$

and recall that this bound holds for all fixed values of $W_1, \dots, W_{k-1}$. So, for any fixed i and fixed $W_1, \dots, W_{k(i,1)-1}$,

$$p_W(\forall j, p_X(E_{k(i,j)}) > \kappa^{-2^{q-i-1}}) < (\kappa^{-2^{q-i-1}})^{\lceil 2^{q-i} \log(q/\varepsilon) \rceil} \leq \kappa^{-\log(q/\varepsilon)/2} < \varepsilon/q,$$

for α chosen large enough. By the union bound, we conclude that with probability at least $1 - \varepsilon$, W satisfies the condition that for each i there is some $j = j'(i)$ such that over the choice of X

$$p_X(E_{k(i,j')}) \leq \kappa^{-2^{q-i-1}}.$$

We claim that this implies that there must be $x \in \mathcal{X}$ with $x \subseteq G(W)$. That is because we can bound

$$\begin{aligned} p_X(|\pi_{k(q, \lceil \log(q/\varepsilon) \rceil)}(X)| > 0) &\leq \sum_{i=1}^q p_X\left(\bigwedge_{j=1}^{\lceil 2^{i-q} \log(q/\varepsilon) \rceil} E_{k(i,j)}\right) \\ &\leq \sum_{i=1}^q p_X(E_{k(i, j'(i))}) \leq \sum_{i=1}^q \kappa^{-2^{q-i-1}} \leq \sum_{i=1}^{\infty} \kappa^{-i/2} < 1, \end{aligned}$$

for α chosen large enough.

It only remains to prove Equation (1). Recall again that the choice of $W_1, \dots, W_{k-1}$ is fixed. Let $E_{k,a}$ denote the following event, which is determined by (W_k, X) :

$$E_{k,a} : \quad |\pi_{k-1}(X)| \leq 2^{q-i+1} \text{ and } |\pi_k(X)| = a.$$

So, $E_k = \bigvee_{a=2^{q-i}}^{2^{q-i+1}} E_{k,a}$. Given a matching z of size a , let $V(z)$ denote the set of $2a$ vertices that are incident to z . For $x \in \mathcal{X}$, say that a pair (W^k, x) , with $|\pi_{k-1}(x)| \leq 2^{q-i+1}$, $|\pi_k(x)| = a$ is *consistent* with a subset $S \subseteq [n]$ and a function $f : [n] \rightarrow [a]$ if the following conditions hold:

1. For all edges $\{u, v\} \in \pi_{k-1}(\mu_k(x))$, either
 - (a) $\{u, v\} \subseteq S$, $\{u, v\} \in \pi_k(x)$ and $f(u) = f(v)$, or
 - (b) $\{u, v\} \cap S = \emptyset$ and $\{u, v\} \notin \pi_k(x)$.
2. If u, v belong to distinct edges of $\pi_k(x)$, then $f(u) \neq f(v)$.

Given any fixed pair $(W_k, x) \in E_{k,a}$, if S and f are chosen uniformly at random, then the probability that the pair is consistent with S, f is at least

$$2^{-2^{q-i+2}} \cdot (1/a^2)^a \cdot a! \geq 2^{-2^{q-i+2}} \cdot (ea)^{-a}.$$

Therefore, if $D_{S,f}$ denotes the event that the pair is consistent with S, f , there is some choice of S, f such that

$$p(E_{k,a} \wedge D_{S,f}) \geq p(E_{k,a}) \cdot 2^{-2^{q-i+2}} \cdot (ea)^{-a}. \quad (2)$$

Fix such a choice of S, f .

With Equation (2) in mind, our goal is to bound $p(E_{k,a} \wedge D_{S,f})$. To do this, we count the number of possible pairs $(W_k, \pi_k(x))$ that can be produced within the event $E_{k,a} \wedge D_{S,f}$. For a matching m , let $V(m)$ denote the vertices incident to m . Given W_k and $\pi_k(x)$, the number of choices for $W_k \cup V(\pi_k(x))$ is

$$\begin{aligned} \binom{n - |W^{k-1}|}{|W_k| + 2a} &\leq \binom{n - |W^{k-1}|}{|W_k|} \cdot \left(\frac{n - |W^{k-1}|}{|W_k|}\right)^{2a} \\ &\leq \binom{n - |W^{k-1}|}{|W_k|} \cdot \left(\frac{2q \cdot \lceil 2^{i-q} \cdot \log(q/\varepsilon) \rceil}{\gamma}\right)^{2a}. \end{aligned} \quad (3)$$

For each choice of $W_k \cup V(\pi_k(x))$, let $x' \in \mathcal{X}$ be the lexicographically first matching satisfying the following conditions:

1. $|\pi_{k-1}(x')| \leq 2^{q-i+1}$, $\pi_{k-1}(x') \subseteq G(W^k \cup V(\pi_k(x)))$,
2. For all $\{u, v\} \in \pi_{k-1}(x')$, either
 - (a) $\{u, v\} \subseteq S$, $f(u) = f(v)$ and $\{u, v\} \subseteq W_k \cup V(\pi_k(x))$, or
 - (b) $\{u, v\} \cap S = \emptyset$.

The conditions given above are satisfied by $\mu_k(x)$, so the choice of x' is well-defined.

Claim 10. $\pi_k(x) \subseteq \pi_{k-1}(x')$.

To prove Claim 10, observe that if $\{u, v\} \in \pi_{k-1}(x')$ and $u, v \in V(\pi_k(x))$, then $u, v \in S$, so the choice of x' implies that $f(u) = f(v)$ and $\{u, v\} \in \pi_k(x)$, because this must be true in $D_{S,f}$. There is no edge $\{u, v\} \in \pi_{k-1}(x')$ such that $u \in V(\pi_k(x))$, $v \notin W^k \cup V(\pi_k(x))$, because such an edge must have $u \in S$, which forces $v \in S$, and hence by 2(a) we get $v \in W_k \cup V(\pi_k(x))$. We conclude that any edge of $\pi_{k-1}(x')$ that is not in $\pi_k(x)$ must be incident to W_k . So, if there is an edge $\{u, v\} \in \pi_k(x)$ that is missing in $\pi_{k-1}(x')$, we must have $\pi_{k-1}(x') \subseteq G(W^k) \cup \pi_k(x)$ and $|\pi_{k-1}(x') \setminus G(W^k)| < |\pi_k(x)|$, but $\mu_k(x)$ is the minimizer, so this cannot happen.

By Claim 10, given $W_k \cup V(\pi_k(x))$ and the event $E_{k,a} \wedge D_{S,f}$, the number of possible choices for $\pi_k(x)$ is bounded by the number of subsets of $\pi_{k-1}(x')$, which is at most $2^{2^{q-i+1}}$. Given $\pi_k(x)$ and $W_k \cup V(\pi_k(x))$, W_k is also determined. Putting these calculations together, and using the spreadness condition of X applied to the matching $\pi_k(x)$, we conclude that for $2^{q-i+1} \geq a \geq 2^{q-i}$,

$$\begin{aligned}
& p(E_{k,a} \wedge D_{S,f}) \\
&= \sum_{(W_k, \pi_k(x)) \in E_{k,a} \wedge D_{S,f}} p(W_k, \pi_k(x)) \\
&= \sum_{(W_k, \pi_k(x)) \in E_{k,a} \wedge D_{S,f}} \binom{n - |W^{k-1}|}{|W_k|}^{-1} \cdot r^{-a} \\
&\leq \binom{n - |W^{k-1}|}{|W_k|} \cdot \left(\frac{2q \cdot \lceil 2^{i-q} \cdot \log(q/\varepsilon) \rceil}{\gamma} \right)^{2a} \cdot 2^{4 \cdot 2^{q-i}} \cdot \binom{n - |W^{k-1}|}{|W_k|}^{-1} \cdot r^{-a} \\
&\leq \left(\frac{2^6 q^2 \cdot \lceil 2^{i-q} \cdot \log(q/\varepsilon) \rceil^2}{r \cdot \gamma^2} \right)^a.
\end{aligned}$$

Using Equation (2) yields the bound

$$p(E_{k,a}) \leq \left(\frac{e 2^8 q^2 \cdot a \cdot \lceil 2^{i-q} \cdot \log(q/\varepsilon) \rceil^2}{r \cdot \gamma^2} \right)^a.$$

Observe that when $2^{i-q} \cdot \log(q/\varepsilon) < 1$, we have

$$a \cdot \lceil 2^{i-q} \cdot \log(q/\varepsilon) \rceil^2 < a,$$

and when $2^{i-q} \cdot \log(q/\varepsilon) \geq 1$, because $a \leq 2^{q-i+1}$, we have

$$\begin{aligned} a \cdot [2^{i-q} \cdot \log(q/\varepsilon)]^2 &\leq 4a \cdot (2^{i-q} \cdot \log(q/\varepsilon))^2 \leq 8 \log^2(q/\varepsilon) = 8(\log(q) + \log(1/\varepsilon))^2 \\ &\leq 16(\log^2(q) + \log^2(1/\varepsilon)). \end{aligned}$$

Recall that $r = \alpha \cdot ((\log t)^2 \cdot (t + \log^2(1/\varepsilon))/\gamma^2)$ and $q = \lceil \log t \rceil$, so in all cases we obtain the bound $p(E_{k,a}) \leq (\alpha/2^{16})^{-a}$. Finally, we bound

$$p(E_k) \leq \sum_{a=2^{q-i}}^{2^{q-i+1}} p(E_{k,a}) \leq (\alpha/2^{17})^{-2^{q-i}},$$

as required.

3. Exponential lower bounds for matching

Here we prove Theorem 3. Let X be a uniformly random matching of size k from $[n/2]$ to $[n] \setminus [n/2]$, and let $\mathcal{X}$ denote its support. Let $Y = G(W)$ (recall that this is the set of edges incident to W), where W is a uniformly random subset of $[n]$ of size q . Note that Y has maximum matching size q . Let $\mathcal{Y}$ denote the support of Y .

If $|m| > k$, then $p(m \subseteq X) = 0$, and if $|m| \leq k$, then by the arithmetic-mean geometric-mean inequality,

$$\begin{aligned} p(m \subseteq X) &= \left(\frac{\binom{n/2}{k}}{\binom{n/2-|m|}{k-|m|}} \right)^{-2} \frac{1}{(k)(k-1) \cdots (k-|m|+1)} \\ &\leq \left(\frac{n}{2k} \right)^{-2|m|} \cdot (k-|m|/2+1/2)^{-|m|} \\ &\leq \left(\frac{n^2}{8k} \right)^{-|m|}, \end{aligned}$$

so X is $(n^2/(8k))$ -spread. For $\eta = 1/(1600\alpha)$, we set

$$\begin{aligned} t &= \min\{q/8, \eta q^2/(k \log^2 n)\}, \quad \varepsilon = \exp(-\eta q/(\sqrt{k} \log n)), \\ r &= \alpha \cdot (\log^2(2t)) \cdot (2t + \log^2(1/\varepsilon)) \cdot (2n/q)^2 \\ &\leq 100\alpha\eta \cdot n^2/k \\ &\leq n^2/(16k). \end{aligned}$$

3.1. Approximation method

Given any set of matchings M , define

$$f_M(x) = \begin{cases} 1 & \text{if } \exists m \in M, m \subseteq x \\ 0 & \text{otherwise.} \end{cases}$$

For each gate g of a given monotone circuit, we compute a set of matchings M_g , such that every $m \in M_g$ is of size at most t . Our approximator for g is f_{M_g} . We define the set of errors:

$$E_g = \{x \in \mathcal{X} : g(x) = 1, f_{M_g}(x) = 0\} \cup \{y \in \mathcal{Y} : g(y) = 0, f_{M_g}(y) = 1\}.$$

We prove that if the circuit has at most $\min\{2^t, 1/\varepsilon\}/2 = \min\{q/8, \eta q/(\sqrt{k} \log n)\}/2$ gates and g is the output gate, then one can compute M_g with

$$p(X \in E_g) \leq 1/2, \text{ and } p(Y \in E_g) \leq 1/2. \quad (4)$$

Given Lemma 5, this is enough to complete the proof of the theorem. Indeed, if the circuit is correct, then when $X \notin E_g$, there must be some matching $Z(X) \in M_g$ such that $Z(X) \subseteq X$, $|Z(X)| \leq t$. For any matching m , we have

$$p(m \subseteq Z(X) | X \notin E_g) \leq \frac{p(m \subseteq Z(X) \subseteq X, X \notin E_g)}{p(X \notin E_g)} \leq 2 \cdot (2r)^{-|m|} \leq r^{-|m|},$$

so $Z(X)$ is r -spread. Lemma 5 then implies that

$$p(g(Y) = 1) \geq p(\exists z \in \mathcal{Z}, z \subseteq Y) - p(Y \in E_g) \geq 1/2 - \varepsilon,$$

proving that the circuit is not correct. A correct circuit must have more than $1/(2\varepsilon)$ gates.

To prove Equation (4), we need to describe how to construct M_g for each gate g . The input gates already correspond to f_{M_g} with M_g containing a single edge, so they are their own approximators. If g is computed from gates h_1, h_2 , we show how to obtain M_g from M_{h_1}, M_{h_2} . If $g = h_1 \vee h_2$, we set $M_g = M_{h_1} \cup M_{h_2}$. This choice guarantees that $E_g \subseteq E_{h_1} \cup E_{h_2}$, so no new errors are introduced.

The more interesting case is when $g = h_1 \wedge h_2$. Then the natural choice is to set

$$M'_g = \{m \cup m' : m \in M_{h_1}, m' \in M_{h_2}, m \cup m' \text{ is a matching}\}.$$

The problem is that M'_g may contain a matching of size $> t$. Given M_g , let

$$N_g = E_g \setminus (E_{h_1} \cup E_{h_2})$$

denote the set of new errors. If $X \in N_g$, then there must be a matching $Z(X) \in M'_g$ with $Z(X) \subseteq X$ and $|Z(X)| \leq 2t$. We construct the final M_g using the following simple algorithm:

Algorithm 1 Computing M_g

$M_g \leftarrow \emptyset$

while $p(X \in N_g) > 0$ **do**

 Find a largest matching m such that $p(m \subseteq Z(X) \mid X \in N_g) \geq r^{-|m|}$.

if $|m| \leq t$ **then**

$M_g \leftarrow M_g \cup \{m\}$

else

halt ; // Terminate if largest valid matching exceeds threshold t

return M_g

At the end of this algorithm, M_g only contains matchings of size at most t , if any. Let us bound the measure of the new errors N_g .

Lemma 11. $p_X(X \in N_g) \leq 2^{-t}$.

Proof. The non-trivial case is when the algorithm discovers a matching m with $|m| > t$. In this case we bound

$$p(X \in N_g) = \frac{p_X(m \subseteq Z, X \in N_g)}{p_X(m \subseteq Z | X \in N_g)} \leq \frac{p_X(m \subseteq X)}{r^{-|m|}} \leq 2^{-|m|} \leq 2^{-t}.$$

□

Lemma 12. $p(Y \in N_g) \leq \varepsilon$.

Proof. If $f_{M_g}(Y) = 1$, let $m(Y) \in M_g$ be the lexicographically first element with $m(Y) \subseteq Y$. Let $V(m(Y))$ denote the set of vertices incident to $m(Y)$. If $p(f_{M_g}(Y) = 1) = 0$, then $p(Y \in N_g) = 0$. Otherwise, we have

$$\begin{aligned} p(Y \in N_g) &\leq p(Y \in N_g \mid f_{M_g}(Y) = 1) \\ &= \sum_{m^*, S \subseteq [n]} p(Y \in N_g \mid m(Y) = m^*, V(m^*) \cap W = S, f_{M_g}(Y) = 1) \\ &\quad \cdot p(m(Y) = m^*, V(m^*) \cap W = S \mid f_{M_g}(Y) = 1). \end{aligned} \tag{5}$$

When m^* was added to M_g , it was chosen as a largest matching satisfying $p(m^* \subseteq Z \mid X \in N_g) \geq r^{-|m^*|}$. Thus, for any matching m' with $V(m') \cap V(m^*) = \emptyset$, we have

$$p(m' \subseteq Z \mid m^* \subseteq Z, X \in N_g) = \frac{p(m' \cup m^* \subseteq Z \mid X \in N_g)}{p(m^* \subseteq Z \mid X \in N_g)} \leq \frac{r^{-(|m^*| + |m'|)}}{r^{-|m^*|}} = r^{-|m'|}.$$

Therefore, $Z \setminus m^*$ conditioned on the event $(m^* \subseteq Z, X \in N_g)$ is an r -spread matching of size at most $2t$. Conditioned on $m(Y) = m^*$ and $V(m^*) \cap W = S$, the random variable Y contains $G(W')$, where W' is a uniformly random subset of the complement of $V(m^*)$ of size

$$|W'| \geq q - |S| \geq q - 4t \geq q/2,$$

since $t \leq q/8$. Applying Lemma 5, we conclude that

$$\begin{aligned} &p(Y \in N_g \mid m(Y) = m^*, V(m^*) \cap W = S, f_{M_g}(Y) = 1) \\ &\leq 1 - p(\exists z \in Z, z \subseteq Y \mid m(Y) = m^*, V(m^*) \cap W = S, f_{M_g}(Y) = 1) \\ &< \varepsilon. \end{aligned}$$

Substituting this bound back into (8) yields

$$\begin{aligned} p(Y \in N_g) &\leq \sum_{m^*, S \subseteq [n]} \varepsilon \cdot p(m(Y) = m^*, V(m^*) \cap W = S \mid f_{M_g}(Y) = 1) \\ &= \varepsilon, \end{aligned}$$

as required. □

Using the union bound with Lemmas 11 and 12, we find that if the circuit has size at most $\min\{2^t, 1/\varepsilon\}/2$, then

$$p\left(X \in \bigcup_g E_g\right) \leq 2^t \cdot 2^{-t}/2 \leq 1/2, \quad p\left(Y \in \bigcup_g E_g\right) \leq (1/(2\varepsilon)) \cdot \varepsilon \leq 1/2,$$

proving Equation (4).

4. Exponential lower bounds for clique

We start by reproducing the ideas of [6] to prove Lemma 2:

Proof of Lemma 2. Without loss of generality, assume that $S = \{n - s + 1, \dots, n\}$. For $i = 1, \dots, n - s$, let W_i be a random subset of $[i]$, where each element is included independently with probability p^{t+s} , sampled independently of Y . Let $W_0 = \emptyset$. Recall that $G(W_i)$ denotes the graph that includes all edges incident to W_i . Let $Y_i = G(W_i) \cup (Y \setminus G([i]))$.

We use a coupling argument to prove that

$$p(\exists x \in \mathcal{X}, \mathcal{C}_{x \cup S} \subseteq Y_i \cup \mathcal{C}_S) \geq p(\exists x \in \mathcal{X}, \mathcal{C}_{x \cup S} \subseteq Y_{i+1} \cup \mathcal{C}_S). \quad (6)$$

Y_0 has the same distribution as Y , and by Lemma 1,

$$p(\exists x \in \mathcal{X}, \mathcal{C}_{x \cup S} \subseteq Y_{n-s} \cup \mathcal{C}_S) > 1 - \varepsilon$$

So, Equation (6) proves Lemma 2.

To prove Equation (6), sample W_i and $T = Y \setminus G([i + 1])$ independently. Then sample $Q \subseteq \{i + 1\}$ by including $i + 1$ with probability p^{t+s} , and sample H as a subset of the edges $\{(i + 1, i + 2), (i + 1, i + 3), \dots, (i + 1, n)\}$ where each edge is included independently with probability p . Set $Y_i = G(W_i) \cup H \cup T$ and set $Y_{i+1} = G(W_i \cup Q) \cup T$.

For each fixing of W_i, T , if there is $x \in \mathcal{X}$ with $\mathcal{C}_{x \cup S} \subseteq G(W_i) \cup T \cup \mathcal{C}_S$, then $\mathcal{C}_{x \cup S} \subseteq Y_i \cup \mathcal{C}_S$ and $\mathcal{C}_{x \cup S} \subseteq Y_{i+1} \cup \mathcal{C}_S$. Similarly, $p(\mathcal{C}_{x \cup S} \subseteq Y_i \cup \mathcal{C}_S | W_i, T) = 0$ if and only if $p(\mathcal{C}_{x \cup S} \subseteq Y_{i+1} \cup \mathcal{C}_S | W_i, T) = 0$. The only interesting case is when both quantities are in $(0, 1)$. In this case, after fixing W_i, T , the probability that H includes the edges required to ensure $\mathcal{C}_{x \cup S} \subseteq G(W_i) \cup H \cup T \cup \mathcal{C}_S$ is at least as large as p^{t+s} , because in the worst case $i + 1$ needs to connect to $t + s$ vertices, and those edges are included in Y_{i+1} with probability at most p^{t+s} . This proves

$$p(\mathcal{C}_{x \cup S} \subseteq Y_i \cup \mathcal{C}_S | W_i, T) \geq p(\mathcal{C}_{x \cup S} \subseteq Y_{i+1} \cup \mathcal{C}_S | W_i, T)$$

and completes the proof of Equation (6). □

Armed with Lemma 2, we can now prove Theorem 8. Let $W \subseteq [n]$ be a uniformly random set of size k , and let $X = \mathcal{C}_W$. Let Y be an Erdos-Renyi random graph where each

edge is included with probability $p = n^{-2/(q-1)}$. We have that W is (n/k) -spread, and the probability that Y includes a clique of size q is at most

$$\binom{n}{q} \cdot p^{\binom{q}{2}} = \binom{n}{q} \cdot n^{-q} \leq 1/2,$$

for $q > 5$. Let α be as in Lemma 2, and set

$$\begin{aligned} t &= \lceil 2(q-1)/\log n \rceil, & \varepsilon &= (n/(8\alpha t k \log n))^{-t} \\ r &= \alpha \log((2t+1)/\varepsilon)/p^{2t+1} \leq 8\alpha t \log n. \end{aligned}$$

4.1. Approximation method

For a family of sets A , define

$$f_A(x) = \begin{cases} 1 & \text{if } \exists S \in A, \exists \mathcal{C}_T \subseteq x \text{ with } S \subseteq T, \\ 0 & \text{otherwise.} \end{cases}$$

Note that we may have $|S| \leq 1$ in the above definition. For each gate g of the circuit, we approximate it with a function f_{A_g} , with $|S| \leq t$ for all $S \in A$. We define the set of errors:

$$E_g = \{x \in \mathcal{X} : g(x) = 1, f_{A_g}(x) = 0\} \cup \{y \in \mathcal{Y} : g(y) = 0, f_{A_g}(y) = 1\}.$$

We prove that if the circuit has at most $\min\{2^t, 1/\varepsilon\}/3$ gates and g is the output gate, then one can compute A_g with

$$p(X \in E_g) \leq 1/3, \text{ and } p(Y \in E_g) \leq 1/3. \quad (7)$$

Given Lemma 2, this is enough to complete the proof of the theorem. Indeed, if the circuit is correct, then when $X \notin E_g$, there must be some set $Z(X) \in A_g$ such that $Z(X) \subseteq W$, $|Z(X)| \leq t$. For any set S , using the spreadness of W , we get

$$p(S \subseteq Z(X) | X \notin E_g) \leq \frac{p(S \subseteq Z(X) \subseteq W, X \notin E_g)}{p(X \notin E_g)} \leq 3 \cdot (n/k)^{-|S|} \leq (n/(3k))^{-|S|},$$

So, $Z(X)$ is r -spread. Lemma 2 implies that

$$p(g(Y) = 1) \geq p(\exists z \in \mathcal{Z}, \mathcal{C}_z \subseteq Y) - p(Y \in E_g) \geq 1 - 1/3 - \varepsilon > 1/2,$$

proving that the circuit is not correct. A correct circuit must have at least $1/(3\varepsilon)$ gates, and that proves Theorem 8.

To prove Equation (7), we need to describe how to construct A_g for each gate g . The input gates already correspond to f_{A_g} with A_g containing a single set of size 2. If g is computed from gates h_1, h_2 , we show how to obtain A_g from A_{h_1}, A_{h_2} . If $g = h_1 \vee h_2$, we set

$A_g = A_{h_1} \cup A_{h_2}$. This choice ensures that $E_g \subseteq E_{h_1} \cup E_{h_2}$, so no new errors are generated. The more interesting case is when $g = h_1 \wedge h_2$. The natural choice is to set

$$A'_g = \{S_1 \cup S_2 : S_1 \in A_{h_1}, S_2 \in A_{h_2}\}.$$

The problem is that A'_g may contain a set of size $> t$. Given A_g , let

$$N_g = E_g \setminus (E_{h_1} \cup E_{h_2})$$

denote the corresponding set of new errors. If $X \in N_g$, there must be a set $Z(X) \in A'_g$ with $Z(X) \subseteq W$ and $|Z(X)| \leq 2t$. We construct the final A_g using the following algorithm:

Algorithm 2 Computing A_g

$A_g \leftarrow \emptyset$

while $p(X \in N_g) > 0$ **do**

 Find a largest set S such that $p(S \subseteq Z(X) \mid X \in N_g) \geq r^{-|S|}$.

if $|S| \leq t$ **then**

$A_g \leftarrow A_g \cup \{S\}$

else

halt ; // Terminate if largest valid set exceeds threshold t

return A_g

At the end of this algorithm, A_g only contains sets of size at most t , if any.

Lemma 13. $p_X(X \in N_g) \leq \varepsilon$.

Proof. The non-trivial case is when the algorithm discovers a set S with $2t \geq |S| > t$. In this case we bound

$$\begin{aligned} p(X \in N_g) &= \frac{p_X(S \subseteq X, X \in N_g)}{p_X(S \subseteq Z \mid X \in N_g)} \leq \frac{p_X(S \subseteq X)}{r^{-|S|}} \leq (n/(8\alpha t k \log n))^{-|S|} \\ &< (n/(8\alpha t k \log n))^{-t}. \end{aligned}$$

□

Lemma 14. $p(Y \in N_g) \leq \varepsilon$.

Proof. If $f_{A_g}(Y) = 1$, let $S(Y) \in A_g$ and $T(Y) \subseteq [n]$ be the lexicographically first pair of sets minimizing $|T(Y)|$ and satisfying $S(Y) \subseteq T$, $\mathcal{C}_{T(Y)} \subseteq Y$. If $p(f_{A_g}(Y) = 1) = 0$, then $p(Y \in N_g) = 0$. Otherwise, we have

$$\begin{aligned} p(Y \in N_g) &\leq p(Y \in N_g \mid f_{A_g}(Y) = 1) \\ &= \sum_{S^*, T^* \subseteq [n]} p(Y \in N_g \mid S(Y) = S^*, T(Y) = T^*, f_{A_g}(Y) = 1) \\ &\quad \cdot p(S(Y) = S^*, T(Y) = T^* \mid f_{A_g}(Y) = 1). \end{aligned} \tag{8}$$

When S^* was added to A_g , it was chosen as a largest matching satisfying

$$p(S^* \subseteq Z \mid X \in N_g) \geq r^{-|S^*|}.$$

Thus, for any set S' with $S^* \cap S' = \emptyset$, we have

$$p(S' \subseteq Z \mid S^* \subseteq Z, X \in N_g) = \frac{p(S' \cup S^* \subseteq Z \mid X \in N_g)}{p(S^* \subseteq Z \mid X \in N_g)} \leq \frac{r^{-(|S^*|+|S'|)}}{r^{-|S^*|}} = r^{-|S'|}.$$

Therefore, $Z \setminus S^* \mid (S^* \subseteq Z, X \in N_g)$ is a random r -spread set of size at most $2t$. Conditioned on $S(Y) = S^*, T(Y) = T^*$, the random variable Y is the union of an Erdos-Renyi graph with $\mathcal{C}_{T^*}$. Applying Lemma 2, we conclude that

$$p(Y \in N_g \mid S(Y) = S^*, T(Y) = T^*, f_{A_g}(Y) = 1) \leq 1 - p(\exists z \in \mathcal{Z}, \mathcal{C}_{S^* \cup z} \subseteq Y \cup \mathcal{C}_{S^*}) < \varepsilon.$$

Substituting this bound back into (8) yields

$$p(Y \in N_g) \leq \sum_{S^*, T^*} \varepsilon \cdot p(S(Y) = S^*, T(Y) = T^*, f_{A_g}(Y) = 1) = \varepsilon.$$

□

Using the union bound with Lemmas 13 and 14, we find that if the circuit has size at most $1/(3\varepsilon)$, then

$$p\left(X \in \bigcup_g E_g\right) \leq (1/(3\varepsilon)) \cdot \varepsilon \leq 1/3, \quad p\left(Y \in \bigcup_g E_g\right) \leq (1/(3\varepsilon)) \cdot \varepsilon \leq 1/3,$$

proving Equation (7).

5. Exponential lower bounds for an explicit function

In this section we prove Theorem 9. Let $w \in D_d$ be uniformly random, and let $X = w([k])$. Let $Y \subseteq [n]$ be uniformly random. We have

$$\begin{aligned} p(|w([k])| \leq k/2) &\leq p(|\{i, j \in [k], i \neq j, w(i) = w(j)\}| \geq k/2) \\ &\leq \frac{\mathbb{E}_w[|\{i, j \in [k], i \neq j, w(i) = w(j)\}|]}{k/2} \\ &\leq \frac{\binom{k}{2}/n}{k/2} \\ &\leq k/n, \end{aligned}$$

So, $p(F_{k,d}(X) = 1) \geq 1 - p(|w([k])| \leq k/2) \geq 1 - k/n \geq 2/3$.

$$p(F(Y) = 1) \leq \sum_{\substack{w \in D_d \\ |w([k])| \geq k/2}} 2^{-k/2} \leq n^{d+1} \cdot 2^{-k/2} \leq 1/3,$$

because $3d \log n < k$. By the d -wise independence of the evaluations of w , for a set S with $|S| \leq d$, we have

$$p(S \subseteq X) \leq p(S \subseteq w([k])) \leq \binom{k}{|S|} \cdot \frac{|S|!}{n^{|S|}} \leq \left(\frac{n}{k}\right)^{-|S|}, \quad (9)$$

so X is (n/k) -spread, at least for sets of size at most d . Set $\beta = 32\alpha$ and

$$\begin{aligned} t &= d/2, & \varepsilon &= \exp(-n/(32\alpha k)) \\ r &= \alpha \log(t/\varepsilon)/(1/2)^2 \leq n/(8k) + \alpha \log d \leq n/(4k). \end{aligned}$$

5.1. Approximation method

For a family of sets A , define

$$f_A(x) = \begin{cases} 1 & \text{if } \exists S \in A, S \subseteq x, \\ 0 & \text{otherwise.} \end{cases}$$

For each gate g of the circuit, we approximate it with a function f_{A_g} , with $|S| \leq t$ for all $S \in A$. We define the set of errors:

$$E_g = \{x \in \mathcal{X} : g(x) = 1, f_{A_g}(x) = 0\} \cup \{y \in \mathcal{Y} : g(y) = 0, f_{A_g}(y) = 1\}.$$

We prove that if the circuit has at most $\min\{2^t, 1/\varepsilon\}/2$ gates and g is the output gate, then one can compute A_g with

$$p(X \in E_g) \leq 1/2, \text{ and } p(Y \in E_g) \leq 1/2. \quad (10)$$

Given Lemma 1, this is enough to complete the proof of the theorem. Indeed, if the circuit is correct, then when $X \notin E_g$, there must be some set $Z(X) \in A_g$ such that $Z(X) \subseteq X$, $|Z(X)| \leq t$. Because $2t \leq d$, by Equation (9), for any set S

$$p(S \subseteq Z(X) | X \notin E_g) \leq \frac{p(S \subseteq Z(X) \subseteq X, X \notin E_g)}{p(X \notin E_g)} \leq 2 \cdot (n/k)^{-|S|} \leq (n/(2k))^{-|S|},$$

So, $Z(X)$ is r -spread. Lemma 1 then implies that

$$p(F(Y) = 1) \geq p(\exists z \in \mathcal{Z}, z \subseteq Y) - p(Y \in E_g) \geq 1 - 1/2 - \varepsilon > 1/3,$$

which is a contradiction. Thus, a correct circuit must have more than $\min\{2^t, 1/\varepsilon\}/2$ gates, proving Theorem 9.

When S^* was added to A_g , it was chosen as a largest matching satisfying

$$p(S^* \subseteq Z \mid X \in N_g) \geq r^{-|S^*|},$$

so for any set S' with $S^* \cap S' = \emptyset$, we have

$$p(S' \subseteq Z \mid S^* \subseteq Z, X \in N_g) = \frac{p(S' \cup S^* \subseteq Z \mid X \in N_g)}{p(S^* \subseteq Z \mid X \in N_g)} \leq \frac{r^{-(|S^*|+|S'|)}}{r^{-|S^*|}} = r^{-|S'|}.$$

Therefore, $Z \setminus S^* \mid (S^* \subseteq Z, X \in N_g)$ is an r -spread set of size at most $2t$. Conditioned on $S(Y) = S^*$, the random variable $Y \setminus S^*$ is a uniformly random subset of the complement of S^* . Applying Lemma 1, we conclude that

$$p(Y \in N_g \mid S(Y) = S^*, f_{A_g}(Y) = 1) \leq 1 - p(\exists z \in \mathcal{Z}, z \subseteq Y \cup S^*) < \varepsilon.$$

Substituting this bound back into (11) yields

$$p(Y \in N_g) \leq \sum_{S^*} \varepsilon \cdot p(S(Y) = S^* \mid f_{A_g}(Y) = 1) = \varepsilon,$$

which completes the proof. $\square$

Using the union bound with Lemmas 15 and 16, we find that if the circuit has size at most $\min\{2^{-t}, 1/\varepsilon\}/2$, then

$$p\left(X \in \bigcup_g E_g\right) \leq 2^t \cdot 2^{-t}/2 \leq 1/2, \quad p\left(Y \in \bigcup_g E_g\right) \leq (1/(2\varepsilon)) \cdot \varepsilon \leq 1/2,$$

proving Equation (7).

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