# Size Bounds on Low Depth Circuits for Promise Majority

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—— **Abstract** ——

We give two results on the size of AC0 circuits computing promise majority. $\epsilon$-promise majority is majority promised that either at most an $\epsilon$ fraction of the input bits are 1, or at most $\epsilon$ are 0.

- First, we show super quadratic lower bounds on both monotone and general depth 3 circuits for promise majority.

  - For any $\epsilon \in (0, 1/2)$, monotone depth 3 AC0 circuits for $\epsilon$-promise majority have size

    $$\tilde{\Omega}\left(\epsilon^3 n^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon)}}\right)$$

  - For any $\epsilon \in (0, 1/2)$, general depth 3 AC0 circuits for $\epsilon$-promise majority have size

    $$\tilde{\Omega}\left(\epsilon^3 n^{2 + \frac{\ln(1-\epsilon^2)}{2\ln(\epsilon)}}\right)$$

  These are the first nontrivial size lower bounds on depth 3 promise majority circuits for $\epsilon < 0.45$.

- Second, we give both uniform and non-uniform sub-quadratic size constant depth circuits for promise majority.

  - For integer $k \geq 1$, constant $\epsilon \in (0, 1/2)$, there exists monotone *non* uniform AC0 circuits of depth $2 + 2 \cdot k$ computing $\epsilon$-promise majority with size

    $$\tilde{O}\left(n^{\frac{1}{1-2^{-k}}}\right)$$

  - For integer $k \geq 1$, constant $\epsilon \in (0, 1/2)$, there exists monotone *uniform* AC0 circuit of depth $2 + 2 \cdot k$ computing $\epsilon$-promise majority with size

    $$n^{\frac{1}{1-\left(\frac{2}{3}\right)^k} + o(1)}$$

  These circuits are based on incremental improvements to existing depth 3 circuits for promise majority given by Ajtai [2] and Viola [16] combined with a divide and conquer strategy.

## 1 Introduction

The majority function is a classic function that cannot be computed in AC0 [9]. But AC0 can compute majority promised the input is either mostly 1s or mostly 0s.

▶ **Definition 1** ($\epsilon$-promise majority)**.** *Let* $W : \{0,1\}^n \to [n]$ *be the function giving the number of ones in the input. Let* $\epsilon \in (0, 1/2)$. *Then define the* $\epsilon$ *promise inputs to be:*

$$Maj_\epsilon^0 = \{x \in \{0,1\}^n : W(x) \leq \epsilon n\}$$
$$Maj_\epsilon^1 = \{x \in \{0,1\}^n : W(x) \geq (1-\epsilon)n\}$$
$$Maj_\epsilon = Maj_\epsilon^0 \cup Maj_\epsilon^1$$

*We say that function $f$ solves the $\epsilon$-promise majority[1] problem if:*

$$f(Maj_\epsilon^0) = 0$$

$$f(Maj_\epsilon^1) = 1$$

*That is $f$ computes the majority promised the input is in $Maj_\epsilon$.*

We give size lower bounds to depth 3 alternating circuits computing $\epsilon$-promise majority. Then we give smaller circuits solving promise majority at larger depths.

## 1.1 Motivation

Promise majority is an important tool in derandomizing circuits. We say a circuit implementing a function $f : \{0,1\}^n \times \{0,1\}^m \to \{0,1\}$ is a randomized circuit for $g : \{0,1\}^n \to \{0,1\}$ if $\Pr_{r \in \{0,1\}^m}[f(x,r) = g(x)] \geq 2/3$. We say that a string $r \in \{0,1\}^m$ is a seed for g.

Adleman [1] showed that any randomized $n$ bit circuit has some $O(n)$ fixed choices of seeds that will always give the correct output on the majority of seeds. Taking majority of these outputs gives the right answer. If the circuit has size $O(n)$, this gives an $O(n^2)$ sized deterministic circuit. Unfortunately, $AC0$ cannot compute majority. However, we can get $O(n)$ seeds where 3/5 always give the right answer. Thus we can use a 2/5-promise majority circuit to derandomize $AC0$.

Ajtai [2] gave a depth 3 circuits of size $O\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}}\right)$ solving the $\epsilon$-promise majority problem. Combining this result with Adleman takes a size n, depth d randomized circuit and gives a depth d + 2 size $O\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}}\right)$ deterministic circuit. This is bigger than the ideal $O(n^2)$ circuit Adleman gave in the unbounded depth setting.

This paper gives new size lower bounds for AC0, depth 3 circuits. We show that Adleman's derandomization of AC0 circuits to get $O(n^2)$ size deterministic circuits using promise majority requires a depth increase of 3. We further show that we can achieve $O(n^2)$ size derandomization with a depth 3 increase.

## 1.2 Our Results

For notation, let $\tilde{O}(x)$ indicate order x up to some polylogarithmic factors. That is:

▶ **Definition 2.** $f(n) = \tilde{O}(g(n))$ *if for some integer c,* $f = O(g(n)\ln(n)^c)$.
$f(n) = \tilde{\Omega}(g(n))$ *if for some integer c,* $f = \Omega(g(n)\ln(n)^c)$.

First, we give a lower bound on monotone depth 3 circuits for promise majority. The best known depth 3 circuits are monotone, so this applies to the state of the art.

▶ **Theorem 3.** *For any $\epsilon \in (0, 1/2)$, a* monotone, *depth 3 alternating circuit solving the $\epsilon$-promise majority problem must have $\tilde{\Omega}\left(\epsilon^3 n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)}}\right)$ gates.*

We follow this up with some worse, but still greater than quadratic, lower bounds on any depth 3 AC0 circuit computing promise majority.

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[1] Prior work often called promise majority "approximate majority" [16, 17]. But approximate majority also refers to the standard notion of approximating a boolean function [5]. To avoid confusion, we follow the convention suggested in [12] to refer to the promise problem version of majority as promise majority.

▶ **Theorem 4.** *For any $\epsilon \in (0, 1/2)$, a depth 3 alternating circuit solving the $\epsilon$-promise majority problem must have $\tilde{\Omega}\left(\epsilon^3 n^{2+\frac{\ln(1-\epsilon^2)}{2\ln(\epsilon)}}\right)$ gates.*

Now that we have proven that we cannot construct quadratic size AC0 circuits in depth 3 for promise majority, we show that we can construct them at higher depth. Minor tweaks to Ajtai's promise majority circuit [2] gives us a depth 4 quadratic size, promise majority[2] circuit.

▶ **Theorem 5.** *For constant $\epsilon \in (0, 1/2)$, there exists non uniform, monotone depth 4 alternating circuits solving the $\epsilon$-promise majority problem with size $O\left(n^2\right)$.*

We then show how to solve $\epsilon$-promise majority with even smaller circuits at higher depths using a divide and conquer strategy.

▶ **Theorem 6.** *For constant $\epsilon \in (0, 1/2)$, there exists non uniform, monotone AC0 circuits of depth $2 + 2 \cdot k$ solving the $\epsilon$-promise majority problem with size $\tilde{O}\left(n^{\frac{1}{1-2^{-k}}}\right)$.*

The above circuit is not explicit, or uniform, so we don't know how to construct it efficiently. We also give a P-Uniform circuit for promise majority: a circuit with a polynomial time algorithm to construct it. It uses a slight improvement on Viola's depth 3 promise majority circuit [16] with a divide and conquer strategy.

▶ **Theorem 7.** *For constant $\epsilon \in (0, 1/2)$, there exists P-uniform, monotone AC0 circuits of depth $2 + 2 \cdot k$ solving the $\epsilon$-promise majority problem with size $n^{\frac{1}{1-\left(\frac{2}{3}\right)^k}+o(1)}$.*

This gives depth 6, $o(n^2)$ size, P-uniform monotone promise majority circuits.

▶ **Corollary 8.** *For constant $\epsilon \in (0, 1/2)$, there exists P-uniform monotone AC0 circuits of depth 6 solving the $\epsilon$-promise majority problem with size $n^{\frac{9}{5}+o(1)}$.*

Thus a P-uniform PRG with $O(n)$ seeds for AC0 could derandomize linear sized AC0 randomized circuits to quadratic size with a depth increase of only 5. Finding such PRGs, or even PRGs with polynomially many seeds, is still open. Though work by Dean Doron, Dana Moshkovitz, Justin Oh and David Zuckerman construct nearly quadratic PRGs conditioned on some complexity theoretic assumptions [7].

## 1.3 Related Work

There are already known polynomial size AC0 circuits for promise majority. The first and smallest such circuit is by Ajtai [2] which gives a non uniform, polynomial size, depth 3 AC0 circuit for computing $\epsilon$-promise majority problem.

Ajtai later gave a uniform, even deterministic log time uniform, AC0 circuit for promise majority [3]. But this circuit has a large depth and its construction is complicated. Viola later gave a simpler P-Uniform AC0 circuit for promise majority [16].

Exact threshold functions in AC0 has been studied extensively. Ragde and Wigderson [14] show that for integer $r > 0$ the threshold $\ln(n)^r$ has AC0 circuits with depth $O(r)$ and size $o(n)$. This improves on a result by Ajtai and Ben-Or [4]. Further, Hastad, Wegener, Wurm, and Yi [10] show that polylogarithmic threshold functions have sub polynomial size constant depth circuits.

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[2] It is important that $\epsilon$ stay bounded away from $1/2$ in the upper bounds. The constant in all these upper bounds depends on $\epsilon$ and increases as $\epsilon$ approaches $1/2$.

Approximate majority in AC0 circuits have been investigated by Amano [5], O'Donnel and Whimmer [13]. O'Donnel and Whimmer gave exponential size lower bounds for AC0 circuits computing majority on most inputs, which were shown to be tight by Amano. This is consistent with promise majority results because most inputs are close to balanced, within a $O(1/\sqrt{n})$ factor, but promise majority is only guaranteed to give majority on inputs that are far from balanced.

For $\epsilon = \left(\frac{1}{2} - \frac{1}{\ln(n)^k}\right)$, Viola shows that a randomized, polynomial size AC0 circuit with depth $k+1$ can solve $\epsilon$-promise majority, and that polynomial size deterministic circuits with depth $k+2$ cannot solve $\epsilon$-promise majority. Further there is a depth $k+3$, polynomial size, deterministic circuit for $\epsilon$-promise majority [17].

The same work [17] gave size lower bounds for depth 3 $\epsilon$-promise majority circuits, but the bounds are less than linear for $\epsilon < \frac{1}{2} - \frac{1}{100}$. Closer analysis gives better lower bounds, but I could not get super linear lower bounds for $\epsilon < \frac{1}{2} - \frac{1}{20}$ with this technique.

A later work by Limaya, Srinivasan and Tripathi [12] showed that polynomial, deterministic AC0 circuits with parity gates of depth $k+1$ also cannot solve $\left(\frac{1}{2} - \frac{1}{\ln(n)^k}\right)$-promise majority.

## 2 Proof Ideas

### 2.1 Monotone Depth 3 Circuit Lower Bounds

For depth 3 promise majority circuits, WLOG assume the first level of gates are AND gates[3]. Call the inputs variables, the first level gates clauses, and the second level gates DNFs.

For $\epsilon = \tilde{O}\left(n^{-1/2}\right)$, all our lower bounds hold trivially as we need at least $\epsilon n$ gates. So we will assume otherwise. We do this so that all restrictions will indeed restrict a positive number of items. We generally won't worry about integrality.

To prove lower bounds of a depth 3 circuit, we construct adversarial restrictions that simplify the circuit while setting too few variables to violate the promise. To do this, we use two main tools. The first is a lemma from Viola [16] that we use to remove gates with very small fan in at the first level.

The second is that any collections of large sets over some variables can have a large fraction of sets hit by a small fraction of variables. To do this, just repeatedly select a variable in at least the average number of sets per variable.

First we show DNFs have $\tilde{\Omega}(n^{1+\alpha})$ clauses for some $\alpha > 0$. To do this, we eliminate small clauses with the first idea, then use the second idea to eliminate many clauses with few 0s. This leaves many clauses while eliminating a large fraction of clauses, thus we started with even more clauses.

Then we show the circuit has $\tilde{\Omega}(n^{2+\alpha})$ clauses. First we use the second idea to remove any very large clauses. This lets us fix clauses to 1 without using too many variables. Then using the second idea again, we can hit many DNFs with few clauses. Thus there must be many clauses so we can't hit every DNF with few clauses.

### 2.2 General Depth 3 Circuit Lower Bounds

The proof for non monotone circuits is similar, but with an additional hurdle. In monotone circuits, setting variables to 0 only makes clauses 0. But with negations, we can actually

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[3] Switching the ANDs to ORs and ORs to ANDs in a circuit solving $\epsilon$-promise majority still solves $\epsilon$-promise majority. To see this, observe that flipping all the input bits will flip a $\text{Maj}_\epsilon^1$ input to a $\text{Maj}_\epsilon^0$ input. Then apply De Morgans.

shrink clauses without eliminating them. This is an issue for showing DNFs must be large, but the rest of the argument only needs minor changes.

The solution is to set adversarial bits probabilistically. We independently set each bit to 1 with probability $\epsilon$. With good probability, this will give an input in $\mathrm{Maj}_\epsilon^0$. Some DNFs then must have a good probability of "noticing" and becoming 0.

With high probability, fixing a small fraction of variables according to $D_\epsilon$ will eliminate many clauses. For some $\alpha > 0$, if the DNF is smaller than $n^{1+\alpha}$ this will make it constant. With good probability, setting the rest of the variables gives an input this DNF must "notice" and become 0. Thus for some input both will happen. But then the DNF outputs 0 with few 0 inputs. This can't happen so the DNF must be larger than $n^{1+\alpha}$.

## 2.3 Small Sized Circuits

To get small circuits, first we amplify the $\epsilon$ promise input to a $\frac{1}{\mathrm{polylog}(n)}$ promise input by taking majority over $O(\ln(\ln(n)))$ length walks on an expander graph. Then we separate our input into polynomially small groups and run a $\frac{1}{\ln(n)}$-promise majority on each. This gives a polynomially smaller layer which satisfies just an $\ln(n)$ factor worse promise. Applying this several times computes majority of the promise input.

Ajtai's promise majority strategy gives a quadratic sized $1/\ln(n)$-promise majority circuit. Using this with the divide and conquer strategy above gives non uniform small circuits.

For our uniform circuit, we look at Viola's circuit [16]. It uses a hitting property that requires $n^{3+o(1)}$ many random walks for each of our n bits, requiring an overall size of $n^{4+o(1)}$. We reduce this by showing it suffices to let each bit only range over random walks starting at that bit, giving an $n^{3+o(1)}$ sized $1/\ln(n)$-promise majority circuit.

Applying this improved version of Viola's depth 3 circuit with our divide and conquer strategy gives our uniform small circuits.

## 2.4 Terminology

We will use a biased input in our proofs.

▶ **Definition 9** ($\epsilon$ Biased Input). *For any $\epsilon \in [0,1]$ the $\epsilon$ biased input $D_\epsilon$ is a random variable over $\{0,1\}^n$ where each bit independently is 1 with probability $\epsilon$.*

As with $\mathrm{Maj}_\epsilon^0$ and $\mathrm{Maj}_\epsilon^1$, n in $D_\epsilon$ is implicit. $D_\epsilon$ is related to $\mathrm{Maj}_\epsilon^0$ by a central limit theorem: $\Pr[D_\epsilon \in \mathrm{Maj}_\epsilon] > \frac{1}{3}$ for large enough n.

We will make sub DNFs by only taking some clauses from a larger DNF.

▶ **Definition 10** (Sub DNF). *Let G be a DNF with clauses $C = \{C_i : i \in [k]\}$ so that $G = \bigvee_{i \in [k]} C_i$. Let $\Lambda \subseteq [k]$ and H be a DNF with $H = \bigvee_{i \in \Lambda} C_i$.*
*Then we say that H is sub DNF of G or G has sub DNF F.*

Restrictions fix some bits in the input to a function. We formalize this as a function that takes unrestricted bits as input and outputs the restricted and unrestricted bits together[4].

▶ **Definition 11** (Restriction). *A restriction $\rho$ on n variables of size m is a function $\rho : \{0,1\}^{n-m} \to \{0,1\}^n$ such that for some constant of fixed bit values $c \in \{0,1\}^m$ and permutation of $[n]$ for bit locations, $\pi$, for all $x \in \{0,1\}^{n-m}$ and $i \in n$:*

$$\rho(x)_i = \begin{cases} c_{\pi_i} & \pi_i \leq m \\ x_{\pi_i - m} & \pi_i > m \end{cases}$$

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[4] This is an equivalent, but slightly nonstandard way to define restrictions.

*We say the size of $\rho$ as $|\rho| = m$ and we denote by $f \upharpoonright_\rho = f \circ \rho$.*

When we compose a DNF, F, with a restriction $\rho$, we say $F \upharpoonright_\rho$ is a DNF which is F with variables restricted in $\rho$ set to their restricted value. We simplify such a DNF to remove any clause that has been set to 0. We count the size of a DNF by its number of clauses.

▶ **Definition 12** (DNF Terms). *For a DNF F, the size of F, $|F|$, is the number of clauses in F. Any DNF that is the constant 1 or 0 function has size 0.*
*We say a DNF F has width w if no clause in F has width greater then w.*

## 3    Monotone Depth 3 Circuit Size Lower Bounds

### 3.1    Removing Small Clauses

We use a result from Viola in [16], Lemma 11. Intuitively, this lemma says for a DNF with small width, either there is some setting to a small number of variables that makes it 0, or under a randomized input it is unlikely to be 0.

▶ **Lemma 13.** *Let G be a DNF with a sub DNF F. Assume for some positive integers w and m, F has width at most w and $\Pr[G(D_\epsilon) = 0] \geq e^{-\epsilon^w \cdot m/w^2}$. Then there exists a restriction $\rho$ with $|\rho| \leq m$ such that $F \upharpoonright_\rho = 0$ and $\Pr[G \upharpoonright_\rho (D_\epsilon) = 0] \geq \Pr[G(D_\epsilon) = 0]$*

The proof is in Appendix A. As a corollary, we can can apply small restrictions to eliminate small width clauses.

▶ **Corollary 14.** *Suppose $\epsilon \in (0, 1/2)$, DNF F and constant $\alpha$ has $\Pr[F(D_\epsilon) = 0] = \Omega(1/n^\alpha)$.*
*Then for sufficiently large n and $w = \log_\epsilon \left( \frac{\ln(n)^5}{n\epsilon \ln(\epsilon)^2} \right)$, there is a restriction $\rho$ restricting at most $m = \frac{\epsilon n}{\ln(n)}$ variables so that any clause C in F with width w has $C \upharpoonright_\rho = 0$ and $\Pr[F \upharpoonright_\rho (D_\epsilon) = 0] \geq \Pr[F(D_\epsilon) = 0]$.*

**Proof.** Let $F'$ be the sub DNF of F with clauses of width less than w. Then see that

$$\mathbb{E}[F(D_\epsilon) = 0] \geq 1/n^\alpha = e^{-\alpha \ln(n)} = e^{-\alpha \frac{\ln(n)^5}{n} \frac{\epsilon n}{\ln(n)} \frac{\ln(\epsilon)^2}{\ln(n)^3}} = e^{-\alpha \epsilon^w m \frac{1}{\ln(n)^3}} \geq e^{-\epsilon^w m \frac{1}{w^2}}$$

From Lemma 13, there is a restriction $\rho$ of size m with $\mathbb{E}[F \upharpoonright_\rho (D_\epsilon) = 0] \geq \mathbb{E}[F(D_\epsilon) = 0]$ setting $F' \upharpoonright_\rho = 0$. Any width w clause C would be in $F'$, thus $C \upharpoonright_\rho = 0$ since $F' \upharpoonright_\rho = 0$. ◀

### 3.2    Covering Many Large Sets with Few Elements

We prove the simplest version of this result, but slight variations will be used in multiple places. Since the proofs look very similar, we only present one in detail. We show how to remove many clauses from a monotone DNF with a small restriction

▶ **Lemma 15.** *Let F be a monotone, width w DNF. Then for any positive integer b, there is some restriction $\rho$ with $|\rho| = b$ only fixing variables to 0 such that $|F \upharpoonright_\rho| < |F|e^{w \ln\left(1 - \frac{b}{n+1}\right)}$*

**Proof.** The idea is to restrict the variable that intersects the most clauses to 0. This removes at least the average number of clauses per variable, which when we have m clauses and have fixed i variables is at least $\frac{mw}{n-i}$. After b restrictions, we get $\rho$ with $|\rho| = b$ and

$$|F \upharpoonright_\rho| \leq |F| \prod_{i=0}^{b-1} \left( 1 - \frac{w}{n-i} \right)$$

We prove this by induction then simplify with calculus. For the base case where $b = 0$, F is unchanged and we get the empty product, so done.

For $b > 0$, we have some $\rho'$ restricting $b - 1$ variables with $|F \restriction_{\rho'}| \leq |F| \prod_{i=0}^{b-2} \left(1 - \frac{w}{n-i}\right)$. Then $F \restriction_{\rho'}$ is a function on $n + 1 - b$ variables. Let $s$ be the variable in the most clauses of $F \restriction_{\rho'}$. Then $s$ is in at least $|F \restriction_{\rho'}| \frac{w}{n+1-b}$ clauses. Let $\rho$ be $\rho'$ also fixing $s$ to 0. Then:

$$|F \restriction_{\rho}| \leq |F \restriction_{\rho'}| - |F \restriction_{\rho'}| \frac{w}{n+1-b} = |F| \prod_{i=0}^{b-1} \left(1 - \frac{w}{n-i}\right)$$

Completing our induction. This simplifies to

$$|F \restriction_{\rho}| = |F| \prod_{i=0}^{b-1} \left(1 - \frac{w}{n-i}\right) < |F| e^{\sum_{i=0}^{b-1} -\frac{w}{n-i}} = |F| e^{-w \sum_{i=n+1-b}^{n} \frac{1}{i}}$$

From calculus we have

$$\sum_{i=a}^{b} \frac{1}{i} \geq \int_{a}^{b+1} \frac{1}{x} dx = \ln\left(\frac{b+1}{a}\right)$$

Then applying this we get

$$|F \restriction_{\rho}| < |F| e^{-w \sum_{i=n+1-b}^{n} \frac{1}{i}} \leq |F| e^{-w \ln\left(\frac{n+1}{n+1-b}\right)} = |F| e^{w \ln\left(1 - \frac{b}{n+1}\right)}$$

◀

The same idea gives the simpler bound:

▶ **Corollary 16.** *Let F be a monotone DNF with clause width at least $w$. Then there is some restriction $\rho$ that fixes $b$ variables such that $|F \restriction_{\rho}| < |F| e^{-wb/n}$.*

With this idea, we can remove all large clauses fixing few variable. For the non monotone case, we only remove half the average number of clauses with each variable, giving:

▶ **Corollary 17.** *Let F be a collection clauses. Then there is some restriction $\rho$ fixing $n/p$ variables such that $F \restriction_{\rho}$ has width $w = 2 \ln(|F|)p$.*

## 3.3 Monotone DNF Size

We prove that any DNF with a good chance of "noticing" inputs from $D_{\epsilon}$ have a large size.

▶ **Lemma 18.** *For $\epsilon \in (0, 1/2)$ and monotone DNF F with $F(Maj_{\epsilon}^{1}) = 1$ and $\Pr[F(D_{\epsilon}) = 0] \geq 1/n^{\alpha}$ for constant $\alpha$. Then F has $\tilde{\Omega}\left(\epsilon n^{1 + \frac{\ln(1-\epsilon)}{\ln(\epsilon)}}\right)$ clauses.*

**Proof.** The idea is to restrict our function until we are only promised it outputs 1 on an $Maj_{\epsilon/\ln(n)}^{1}$ input. Using Lemma 15, we can do this in such a way to eliminate a large fraction of clauses. Then since we still need to output 1 if we have fewer than $\frac{\epsilon n}{\ln(n)}$ more 0s, we can choose these remaining $\frac{\epsilon n}{\ln(n)}$ 0s to each eliminate one clause, showing that there are still $\frac{\epsilon n}{\ln(n)}$ clauses left. This implies that we must have started with the above number of clauses.

For $w = \log_{\epsilon}(\frac{\ln(n)^5}{n\epsilon \ln(\epsilon)^2})$, by Corollary 14, there is restriction $\rho$ with $|\rho| \leq \frac{\epsilon n}{\ln(n)}$ and $F \restriction_{\rho}$ has no clauses smaller than w. Denote $F_2 = F \restriction_{\rho}$. Note $F_2$ solves $F_2(Maj_{\epsilon(1-1/\ln(n))}^{1}) = 1$ and has no clauses smaller than $w$.

Now use Lemma 15 to get restriction $\rho_2$ that assigns $\epsilon n(1 - 2/\ln(n))$ variables and:

$$|F_2 \restriction_{\rho_2}| \leq |F_2| e^{w \ln\left(1 - \frac{\epsilon n(1 - 2/\ln(n))}{n+1}\right)}$$

Now we simplify the exponent $w \ln\left(1 - \frac{\epsilon n(1 - 2/\ln(n))}{n+1}\right)$. For $0 < x < 1/2$ and $0 < y$ by a Taylor argument we have $\ln(1 - x + y) \leq \ln(1 - x) + 2y$. Then for sufficiently large n:

$$\ln\left(1 - \frac{\epsilon n(1 - 2/\ln(n))}{n+1}\right) = \ln\left(1 - \epsilon + \frac{\epsilon}{n+1} + \frac{2\epsilon n}{n+1}\frac{1}{\ln(n)}\right) \leq \ln(1 - \epsilon) + \frac{5\epsilon}{\ln(n)}$$

Now including w

$$w \ln\left(1 - \frac{\epsilon n(1 - 2/\ln(n))}{n+1}\right) = \frac{\ln(n) - \ln\left(\frac{\ln(n)^5}{\epsilon \ln(\epsilon)^2}\right)}{\ln(1/\epsilon)}\left(\ln(1 - \epsilon) + \frac{5\epsilon}{\ln(n)}\right)$$

$$< \frac{\ln(n)\ln(1 - \epsilon)}{\ln(1/\epsilon)} - \frac{\ln\left(\frac{\ln(n)^5}{\epsilon \ln(\epsilon)^2}\right)\ln(1 - \epsilon)}{\ln(1/\epsilon)} + \frac{5}{\ln(1/\epsilon)}$$

Then applying this to our size bound

$$|F_2 \restriction_{\rho_2}| \leq |F_2| e^{w \ln\left(1 - \frac{\epsilon n(1 - 2/\ln(n))}{n+1}\right)}$$

$$< |F_2| e^{\frac{\ln(n)\ln(1 - \epsilon)}{\ln(1/\epsilon)} - \frac{\ln\left(\frac{\ln(n)^5}{\epsilon \ln(\epsilon)^2}\right)\ln(1 - \epsilon)}{\ln(1/\epsilon)} + \frac{5}{\ln(1/\epsilon)}}$$

$$< |F_2| e^8 n^{\frac{\ln(1 - \epsilon)}{\ln(1/\epsilon)}} \frac{\ln(n)^5}{(1 - \epsilon)\ln(\epsilon)^2 \log_\epsilon(1 - \epsilon)}$$

$$< |F_2| e^8 n^{\frac{\ln(1 - \epsilon)}{\ln(1/\epsilon)}} 2\ln(n)^5$$

Since $F_2 \restriction_{\rho_2}$ still satisfies $F_2 \restriction_{\rho_2} (\mathrm{Maj}^1_{\epsilon/\ln(n)}) = 1$, it must have $\frac{\epsilon n}{\ln(n)}$ more clauses. Thus:

$$\frac{\epsilon n}{\ln(n)} \leq |F_2 \restriction_{\rho_2}| \leq e^8 |F_2| n^{\frac{\ln(1 - \epsilon)}{\ln(1/\epsilon)}} 2\ln(n)^5$$

$$\frac{\epsilon n^{1 + \frac{\ln(1 - \epsilon)}{\ln(\epsilon)}}}{2e^8 \ln(n)^6} \leq |F_2|$$

$F$ has at least as many clauses as $F_2$, thus $|F| = \tilde{\Omega}\left(\epsilon n^{1 + \frac{\ln(1 - \epsilon)}{\ln(\epsilon)}}\right)$. ◀

## 3.4 Monotone Circuit Size Lower Bounds

Now we prove the monotone depth 3 promise majority circuit lower bounds.

▶ **Theorem 3.** *For any $\epsilon \in (0, 1/2)$, a* monotone, *depth 3 alternating circuit solving the $\epsilon$-promise majority problem must have* $\tilde{\Omega}\left(\epsilon^3 n^{2 + \frac{\ln(1 - \epsilon)}{\ln(\epsilon)}}\right)$ *gates.*

**Proof.** Let $F$ be a monotone depth 3 alternating circuit computing the $\epsilon$-promise majority. We will refer to the first level gates as clauses, and the second level gates as DNFs. Let $|F|$ refer to the number of clauses in F, and $\|F\|$ refer to the number of DNFs. If F has more than $n^{2+\alpha}$ gates, we are done. So suppose it doesn't.

Let $\alpha = \frac{\ln(1 - \epsilon + 3\epsilon/\ln(n))}{\ln(\epsilon - 3\epsilon/\ln(n))}$. We can show that

$$\alpha > \frac{\ln(1 - \epsilon)}{\ln(\epsilon)} - O\left(\frac{1}{\ln(n)}\right)$$

So if we show $|F| = \tilde{\Omega}\left(\epsilon^3 n^{2+\alpha}\right)$, then the O term in $\alpha$ becomes a constant.

First, from Corollary 17, we have a restriction $\rho$ fixing $\epsilon n/\ln(n)$ variables such that any clause wider than $w = 2\ln(|F|)\frac{\ln(n)}{\epsilon}$ is set to 0. Let $F_2 = F \restriction_\rho$. See that $F_2$ solves the $\left(\epsilon - \frac{1}{\ln(n)}\right)$-promise majority problem and has no clauses wider than $6\frac{\ln(n)^2}{\epsilon}$.

By Lemma 18, every DNF $G$ with $\Pr[G(D_{\epsilon(1-3/\ln(n))}) = 0] \geq 1/n^{3+\alpha}$ has at least $c\epsilon n^{1+\alpha}$ clauses for some polylogarithm $c$. Let $F_3$ be the sub circuit of $F_2$ with only the DNFs of $F_2$ larger than $c\epsilon n^{1+\alpha}$.

Since no clauses are wider than $w$, we can set any $m$ clauses in $F_3$ to 0 by fixing only $mw$ variables. Then, analogous to Corollary 16, there exists a restriction $\rho_2$ fixing $\epsilon n/\ln(n)$ variables to one such that:

$$\|F_3 \restriction_{\rho_2}\| \leq \|F_3\| e^{-c\epsilon n^{1+\alpha}(|\rho_2|/w)/|F_3|}$$

Where $\|F_3 \restriction_{\rho_2}\|$ is the number of DNFs in $F_3$ not fixed to 1 or 0 under the restriction $\rho_2$.

See that $F_2 \restriction_{\rho_2}$ still solves the $\epsilon(1 - 2/\ln(n))$-majority problem. By a central limit theorem $D_{\epsilon(1-3/\ln(n))}$ has a constant above 0 probability of being in $\mathrm{Maj}^0_{\epsilon(1-3/\ln(n))}$. Thus since $F_2$ has fewer than $n^{2+\alpha}$ DNFs (by assumption), some DNF in $F_2$, A, must be 0 on $D_{\epsilon(1-3/\ln(n))}$ with probability greater than $1/n^{3+\alpha}$. By Lemma 18, A has size at least $c\epsilon n^{1+\alpha}$. Thus A must also be in $F_3$. Thus $\|F_3 \restriction_{\rho_2}\| \geq 1$.

Now we can compute a lower bound for $|F_3|$:

$$1 \leq \|F_3 \restriction_{\rho_2}\| \leq \|F_3\| e^{-c\epsilon n^{1+\alpha}\frac{|\rho_2|}{w|F_3|}}$$

$$e^{c\epsilon^2 n^{2+\alpha}\frac{1}{w|F_3|\ln(n)}} \leq \|F_3\|$$

$$c\epsilon^2 n^{2+\alpha}\frac{1}{w|F_3|\ln(n)} \leq \ln\left(\|F_3\|\right)$$

$$c\epsilon^3 n^{2+\alpha}\frac{1}{2\ln(|F|)\ln(n)|F_3|\ln(n)} \leq \ln\left(\|F_3\|\right)$$

$$\tilde{\Omega}\left(\epsilon^3 n^{2+\alpha}\right) \leq |F_3|$$

Using the definition of alpha and that $|F| > |F_3|$ we get:

$$|F| \geq \tilde{\Omega}\left(\epsilon^3 n^{2+\alpha}\right) \geq \tilde{\Omega}\left(\epsilon^3 n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)}-O\left(\frac{1}{\ln(n)}\right)}\right) \geq \tilde{\Omega}\left(\epsilon^3 n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)}}\right)$$

◀

## 4    General Depth 3 Circuits

The proof the size lower bound for general depth 3 circuits computing promise majority follows almost the same as the monotone case, except the proof that DNFs must be large. So we only prove our DNF bound here, but leave a brief proof of Theorem 4 in Appendix B.

▶ **Lemma 19.** *Suppose $\epsilon \in (0, 1/2)$, and $F$ is a DNF such that $\Pr[F(D_\epsilon) = 0] \geq 1/n^\alpha$ for some constant $\alpha$ and $F(\mathrm{Maj}^1_\epsilon) = 1$. Then $F$ has size at least $\tilde{\Omega}\left(\epsilon n^{1+\frac{\ln(1-\epsilon^2)}{2\ln(\epsilon)}}\right)$.*

**Proof.** First, see that if $F(\mathrm{Maj}^1_\epsilon) = 1$ and $F \neq 1$, there must be at least $\epsilon n$ clauses. Otherwise we could fix one variable in each clause to 0 using fewer than $n\epsilon$ 0s. Then for $\epsilon = \tilde{O}\left(\frac{1}{\sqrt{n}}\right)$ the lemma is satisfied. So take $\epsilon = \omega\left(\frac{\ln(n)^3}{\sqrt{n}}\right)$.

Let $m = \epsilon n(1 - 2/\ln(n))$ and $w = \log_\epsilon\left(\frac{\ln(n)^5}{n\epsilon \ln(\epsilon)^2}\right)$. We will define a sequence of probabilistic restrictions, $\rho_0, ..., \rho_m$, each restricting one more variable according to $D_\epsilon$. At the same time

we will construct a sequence of sub DNFs of F, $F_0, ..., F_m$, each a subset of the last, so that each $F_i \upharpoonright_{\rho_i}$ has width at least w.

Informally, with decent probability each $F_i$ is significantly smaller than the last. Thus by a Chernoff bound, with high probability $F_m$ has a small fraction of the clauses of F. Then we use Corollary 14 to eliminate the small width clauses in $F_m \upharpoonright_{\rho_m}$. With good probability the DNF will still not be 1, in which case it must still have an almost linear number of clauses. Thus there must have been many clauses to destroy so many and have so many left.

Let $\rho_0$ restrict no variables and $F_0$ be F restricted to clauses wider than w. Then for any i, let $\rho_i$ be $\rho_{i-1}$ plus restricting whichever variable appears in the most clauses in $F_{i-1} \upharpoonright_{\rho_{i-1}}$ to one with probability $\epsilon$ and 0 otherwise. Then let $F_i$ be the clauses such that they have width higher than w in $F \upharpoonright_{\rho_i}$. See that $F_i \subseteq F_{i-1}$, since further restrictions will only decrease the size and number of clauses.

With probability at least $\epsilon$, $\rho_i$ will eliminate at least $\frac{|F_{i-1}|w}{2(n-i+1)}$ clauses. Thus:

$$\Pr\left[|F_{i+1}| \leq |F_i|\left(1 - \frac{w}{2(n-i)}\right)\right] \geq \epsilon$$

Let k be the number of times this inequality holds. By a argument similar to Lemma 15:

$$|F_m| \leq |F_0| \prod_{i=0}^{k-1}\left(1 - \frac{w}{2(n-i)}\right)^k \leq e^{\frac{w}{2}\ln\left(1 - \frac{k}{n+1}\right)}$$

See the expected value of k is at least $m\epsilon$. By a Chernoff bound, we have:

$$\Pr[k < (1 - 1/\ln(n))\epsilon m] \leq e^{-\frac{\epsilon m}{2\ln(n)^2}} < e^{-\frac{\epsilon^2 n}{\ln(n)^3}}$$

Now, notice that $\rho_m$ only sets variables according to an $\epsilon$ biased distribution. So if we just finish sampling the rest of the variables from $D_\epsilon$ it is the same as sampling all the variables from $D_\epsilon$. Thus:

$$\mathbb{E}_{\rho_m}[\Pr[F \upharpoonright_{\rho_m} (D_\epsilon) = 0]] = \Pr[F(D_\epsilon) = 0]$$

We need high probability that $F \upharpoonright_{\rho_m}$ still outputs 0 with polynomial probability on $D_\epsilon$.

$$\frac{1}{n^\alpha} \leq \mathbb{E}_{\rho_m}[\Pr[F \upharpoonright_{\rho_m} (D_\epsilon) = 0]]$$

$$\leq \frac{1}{n^{2\alpha}} + \Pr_{\rho_m}\left[\Pr[F \upharpoonright_{\rho_m} (D_\epsilon) = 0] > \frac{1}{n^{2\alpha}}\right]$$

$$\frac{1}{n^\alpha} - \frac{1}{n^{2\alpha}} \leq \Pr_{\rho_m}\left[\Pr[F \upharpoonright_{\rho_m} (D_\epsilon) = 0] > \frac{1}{n^{2\alpha}}\right]$$

The probability that $\rho_m$ has both $\Pr[F \upharpoonright_{\rho_m} (D_\epsilon) = 0] > 1/n^{2\alpha}$ and $k > (1 - \frac{1}{\ln(n)})\epsilon m$ is at least $\frac{1}{n^\alpha} - \frac{1}{n^{2\alpha}} - e^{-\frac{\epsilon^2 n}{\ln(n)^3}}$, which for large n is positive. Then take such $\rho_m$ as $\rho$.

By Corollary 14, we have a restriction of $F|_\rho$, $\rho'$, which restricts $\epsilon n/\ln(n)$ variables and leaves no clauses of width less than w, and has

$$\Pr[F \upharpoonright_\rho \upharpoonright_{\rho'} (D_\epsilon) = 0] \geq \Pr[F \upharpoonright_\rho (D_\epsilon) = 0] \geq \frac{1}{n^{2\alpha}}$$

Now call $F' = F \upharpoonright_\rho \upharpoonright_{\rho'}$. See that $F'$ has fixed $\epsilon n(1 - 1/\ln(n))$ variables. Thus it still satisfies $F'(\text{Maj}^1_{\epsilon/\ln(n)}) = 1$. Since $F' \neq 1$, $|F'| \geq \epsilon n/\ln(n)$. The clauses in $F'$ had width greater

than w in $F_m$, otherwise $\rho'$ would have set them to 0. Thus $|F_m| \geq \epsilon n / \ln(n)$. Together we have:

$$\frac{\epsilon n}{\ln(n)} \leq |F_0| e^{\frac{w}{2} \ln\left(1 - \frac{k}{n+1}\right)}$$

$$\leq |F_0| e^{\frac{w}{2} \ln\left(1 - \frac{(1 - 1/\ln(n))\epsilon m}{n+1}\right)}$$

$$\leq |F_0| e^{\frac{\ln\left(\frac{n\epsilon \ln(\epsilon)^2}{\ln(n)^5}\right)}{2\ln(1/\epsilon)}\left(\ln\left(1 - \epsilon^2\right) + 6\epsilon/\ln(n)\right)}$$

$$\tilde{\Omega}\left(\epsilon n^{1 + \frac{\ln\left(1 - \epsilon^2\right)}{2\ln(\epsilon)}}\right) \leq |F_0|$$

Thus F has at least $\tilde{\Omega}\left(\epsilon n^{1 + \frac{\ln\left(1 - \epsilon^2\right)}{2\ln(\epsilon)}}\right)$ clauses. ◀

## 5 Circuit Upper Bounds

The results in this section mostly use standard techniques and the details are given in the appendix. A small improvement to Ajtai's [2] promise majority circuit (shown in Appendix C.1) gives:

▶ **Theorem 20.** *For any $\epsilon \in (0, 1/2)$ there exists a monotone depth 3 AC0 circuit solving the $\epsilon$-promise majority problem with size $O\left((\epsilon \ln(\epsilon))^2 n^{2 + \frac{\ln(1 - \epsilon)}{\ln(\epsilon) - \ln(1 - \epsilon)}}\right)$.*

This also gives the corollary we will use for our stronger upper bounds for higher depth.

▶ **Corollary 21.** *For any $\epsilon = O\left(\frac{\ln(\ln(n))}{\ln(n)}\right)$, there is a monotone depth 3 AC0 circuit solving the $\epsilon$-promise majority problem with size $O\left(n^2\right)$.*

Using random walks on expander graphs, we can amplify our promise (proof in Appendix C.2). The polylogarithmic factor in the size depends on $\epsilon$ and k.

▶ **Lemma 22.** *For any constant k, $\epsilon \in (0, 1/2)$, there exists a monotone, P-Uniform depth 3 circuit of size $\tilde{O}(n)$ amplifying a $Maj_\epsilon^0$ input to a $Maj_{\frac{1}{\ln(n)^k}}^0$ output and a $Maj_\epsilon^1$ input to a $Maj_{\frac{1}{\ln(n)^k}}^1$ output.*

With amplification and quadratic size circuits, we can trivially prove the existence of depth 4, $\tilde{O}(n^2)$ sized circuits. But the circuit size only depends on the number of potential inputs (NOT the number of bits used to represent them). Thus the circuit has $O(n^2)$ size.

▶ **Theorem 5.** *For constant $\epsilon \in (0, 1/2)$, there exists non uniform, monotone depth 4 alternating circuits solving the $\epsilon$-promise majority problem with size $O\left(n^2\right)$.*

We can apply promise majority circuits in a divide and conquer fashion to get the following (Proved in Appendix C.3).

▶ **Lemma 23.** *If there is a depth 3 alternating circuit with size $n^\alpha$ solving $1/\ln(n)$-promise majority, then for any positive integer k, there is a depth $1 + 2 \cdot k$ alternating circuit solving $1/\ln(n)^k$-promise majority with size:*

$$kn^{\frac{1}{1 - \left(\frac{\alpha - 1}{\alpha}\right)^k}}$$

*Which is uniform and monotone if the depth 3 circuit is uniform and monotone.*

Combining this with amplification and our quadratic sized majority gives:

▶ **Theorem 6.** *For constant $\epsilon \in (0, 1/2)$, there exists* non *uniform, monotone AC0 circuits of depth $2 + 2 \cdot k$ solving the $\epsilon$-promise majority problem with size $\tilde{O}\left(n^{\frac{1}{1-2^{-k}}}\right)$.*

For uniform circuits, we refine of Viola's result [16].

▶ **Theorem 24.** *There exists a P-uniform, size $O\left(n^{3+o(1)}\right)$ depth 3 AC0 circuit solving the $\frac{1}{\ln(n)}$-promise majority problem.*

The idea of this circuit is that if we get an input in $\mathrm{Maj}^0_{1/2 \ln(n)}$, there is no way to spread $\ln(n)$ copies in a way that hits n bits. But if the input is in $\mathrm{Maj}^1_{1/2 \ln(n)}$ there is a way to spread $\ln(n)$ copies in a way that hits all bits using one of a polynomial number of random walks. We improve the results of Viola by showing a more efficient way to use the same length random walk. Details in Appendix D.1.

Again applying amplification and divide and conquer we get:

▶ **Theorem 7.** *For constant $\epsilon \in (0, 1/2)$, there exists P-uniform, monotone AC0 circuits of depth $2 + 2 \cdot k$ solving the $\epsilon$-promise majority problem with size $n^{\frac{1}{1-\left(\frac{2}{3}\right)^k}+o(1)}$.*

## 6 Closing Statements & Open Problems

For constant $\epsilon$, we have shown that depth 3 circuits solving the $\epsilon$-promise majority problem have greater than quadratic size. For $\epsilon$ bigger than inverse polynomial but smaller than constant, the size of depth 3 $\epsilon$-promise majority must be almost quadratic.

For $\epsilon$ bounded by a constant under $1/2$, we can solve $\epsilon$-promise majority with a constant depth circuit that has size within an arbitrarily small polynomial factor of linear.

But our lower bounds and known circuits are not tight for all choices of $\epsilon$ and not for depth greater than 3. Further its possible that one of our main uses for promise majority, derandomization, may not actually need it. So we still have these open problems:

1. This shows that using Adleman's derandomization of constant depth circuits, or any conventional use of PRGs, requires a depth increase of 3 to have only quadratic overhead. Is there a way to derandomize any constant depth circuit with only depth increase of 2 and quadratic size overhead?
2. Do negations help solve promise majority?
3. What are the optimal depth 3 bounds? For constant $\epsilon$, there is a polynomial gap between our lower bounds $\tilde{\Omega}\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)}}\right)$, and our upper bounds $O\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}}\right)$.
   For $\epsilon \leq \frac{1}{\sqrt{n}}$, optimal circuit size is given trivially by a DNF. But for $\epsilon = n^{-\alpha}$ where $\alpha \in (0, 1/2)$, there is a polynomial gap between our lower bounds $\tilde{\Omega}\left(n^{2-3\alpha}\right)$ and our upper bounds $\tilde{O}\left(n^{2-2\alpha}\right)$.
4. What are the size lower bounds for depth greater than 3?
5. Do these bounds extend to AC0 with parity, or other circuit classes below TC0?

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## A    Proof of Lemma 13

This proof was first presented by Viola [16] in the more straightforward case where $G = F$: where there is only one DNF. We use this lemma on non monotone circuits to eliminate small clauses in a DNF without fixing it to one, so we have to present it in this slight generalization. This does not require any new ideas, just more careful analysis.

▶ **Lemma 13.** *Let G be a DNF with a sub DNF F. Assume for some positive integers w and m, F has width at most w and* $\Pr[G(D_\epsilon) = 0] \geq e^{-\epsilon^w \cdot m/w^2}$. *Then there exists a restriction $\rho$ with $|\rho| \leq m$ such that $F \restriction_\rho = 0$ and $\Pr[G \restriction_\rho (D_\epsilon) = 0] \geq \Pr[G(D_\epsilon) = 0]$*

**Proof.** The proof follows closely from a basic idea: if there are many small independent clauses in a DNF, it will accept with high probability. If there aren't, then we can set a small number of variables to decrease the size of each clause.

See that since F is a sub DNF of G, for any $\rho$, $G \restriction_\rho (x) = 0 \implies F \restriction_\rho (x) = 0$. Thus for any distribution y, $\Pr[F \restriction_\rho (y) = 0] \geq \Pr[G \restriction_\rho (y) = 0]$.

We will use a proof by induction. First the base case, where $w = 1$. If there are only m clauses in F, then let $\rho$ be restriction setting each variable in F to make its corresponding clause 0. Then since this is the only assignment to these variables so that G is not one, we have

$$\Pr[G(D_\epsilon) = 0] \leq \Pr[G \restriction_\rho (D_\epsilon) = 0]$$

Otherwise F must have at least m + 1 clauses. Then $\Pr[F(D_\epsilon) = 0]$ is just the probability that none of its at least m variables is one. Thus:

$$\Pr[G(D_\epsilon) = 0] \leq |\Pr[F(D_\epsilon) = 0] \leq (1 - \epsilon)^{m+1} < e^{-\epsilon m}$$

but by assumption $\Pr[G(D_\epsilon) = 0] \geq e^{-\epsilon m}$, thus this case cannot occur.

Now for the general case where $w \geq 2$. We say that a set of variables, S, is a cover for DNF F if every clause in F contains at least one variable from S. We say that a set of clauses T is independent if no two clauses in T have any variables in common.

Consider a minimum cover of F, S. Let $T$ be a maximal independent set of clauses in F. See that $|T|w \geq |S|$ because if we take every variable in a maximal independent set of clauses, we must get a cover for F (otherwise we would have another independent clause) and we already said S is the smallest cover.

Then the probability the DNF is 0 is at most the probability each clause in $T$ is 0. Thus:

$$
\begin{aligned}
\Pr[F(D_\epsilon) = 0] \leq & (1 - \epsilon^w)^{|T|} \\
< & e^{-|T|\epsilon^w} \\
\leq & e^{-|S|\epsilon^w/w}
\end{aligned}
$$

Now if $|S| \geq m/w$, then we have violated the assumption and are done. Thus we have $|S| < m/w$.

Choose some assignment to the variables in S to get $\rho'$ so that $\Pr[G \restriction_{\rho'} (D_\epsilon) = 0] \geq \Pr[G(D_\epsilon) = 0]$. By an averaging argument, some $\rho'$ must do this. This gives

$$
\begin{aligned}
\Pr[G \restriction_{\rho'} (D_\epsilon) = 0] \geq & \Pr[G(D_\epsilon) = 0] \\
\geq & e^{-\epsilon^w \cdot m/w^2} \\
> & e^{-\epsilon^{w-1} \cdot m \frac{w-1}{w-1} \frac{1}{(w-1)w}} \\
\geq & e^{-\epsilon^{w-1} \cdot m \frac{1-1/w}{(w-1)^2}}
\end{aligned}
$$

Thus $F \restriction_{\rho'}$ is a new width $w - 1$ instance. Then by the inductive hypotheses, we have a restriction $\rho''$ of size at most $m(1 - 1/w)$ setting $F \restriction_{\rho'} \restriction_{\rho''} = 0$ and not decreasing the probability $G \restriction_{\rho'} (D_\epsilon)$ is 0. Thus letting $\rho = \rho' \circ \rho''$, we have

$$|\rho| \leq m/w + m(1 - 1/w) = m$$

$$F \restriction_\rho = 0$$

$$\Pr[G \restriction_\rho (D_\epsilon) = 0] \geq \Pr[G(D_\epsilon) = 0]$$

◀

## B    Proof of Theorem 4

Here We present the proof for the general case depth 3 circuit for promise majority lower bounds. As previously mentioned, the proof is very similar to the proof of Theorem 3.

▶ **Theorem 4.** *For any $\epsilon \in (0, 1/2)$, a depth 3 alternating circuit solving the $\epsilon$-promise majority problem must have $\tilde{\Omega}\left(\epsilon^3 n^{2 + \frac{\ln(1-\epsilon^2)}{2\ln(\epsilon)}}\right)$ gates.*

**Proof.** As before, we first use $\epsilon n / \ln(n)$ variables to remove all very large clauses. Then we will use $\epsilon n / \ln(n)$ more getting rid of many DNFs that output 0 on $D_{\epsilon(1-3/\ln(n))}$ with polynomial probability. This will leave us with almost n DNFs that also must output 0 on $D_{\epsilon(1-3/\ln(n))}$ with high probability.

Let F be a polynomial size, depth 3 circuit computing $\epsilon$-promise majority. Then as in the proof of Theorem 3, assume its first level of gates are AND gates, call them clauses and our second level gates DNFs.

First using Corollary 17, there is a restriction $\rho$ with $|\rho| = \epsilon n / \ln(n)$ variables and removes any clauses width greater than $O(\ln(n)^2 / \epsilon)$. Then let $F' = F|_\rho$.

Now let $F_2$ be the subcircuit of F' that only has DNFs of size at least $c\epsilon n^{1+\alpha}$ for $\alpha = \frac{\ln(1-\epsilon^2)}{2\ln(\epsilon)}$ and some $c = \mathrm{polylog}(n)$ given by Lemma 19. Now we fix $\epsilon^2 n / \ln(n)^4$ clauses in $F_2$ that eliminate as many DNFs as we can to get $F_3$. This requires us to fix fewer than $\epsilon n / \ln(n)$ variables. By a similar proof to Theorem 3, we can get

$$\|F_3\| \le \|F_2\| e^{-\epsilon^3 \frac{n^{2+\alpha}}{|F_2|\mathrm{polylog}(n)}}$$

Now, $\|F_3\|$ must have at least 1 DNF of size $c\epsilon n^{1+\alpha}$ because $F_3$ still solves a $\epsilon\left(1 - \frac{2}{\ln(n)}\right)$-promise majority problems. Thus:

$$1 \le \|F_2\| e^{-\epsilon^3 \frac{n^{2+\alpha}}{|F_2|\mathrm{polylog}(n)}}$$

This only holds if $|F_2| = \tilde{\Omega}(\epsilon^3 n^{2+\alpha})$ or $\|F_2\|$ is exponentially large. Thus the size of the circuit is $\tilde{\Omega}(\epsilon^3 n^{2+\alpha})$.    ◀

## C    Circuit Upper Bounds

### C.1    Ajtai's Construction

Ajtai gave a randomized construction of a depth 3 circuit for promise majority. It is defined recursively with each layer being a conjunction or disjunction of a specific number of random circuits of smaller depth.

Intuitively, it gives a circuit which solves a coin problem, determining whether an input comes from $D_\epsilon$ or $D_{1-\epsilon}$, with exponentially high probability. Then feeding in random indexes from a $\mathrm{Maj}_\epsilon$ instance is equivalent to feeding in an input from $D_\epsilon$ or $D_{1-\epsilon}$. Then by an averaging argument, some choice of indexes must always work.

So first we prove:

▶ **Lemma 25.** *For any $\epsilon \in (0, 1/2)$ and positive integer n, there exists a monotone depth 3 alternating circuit, C, with size $O\left(n^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}}\right)$ so that for any $\delta \le \epsilon$:*

$$\Pr[C(D_\delta) = 1] < 2^{-n}$$

$$\Pr[C(D_{1-\delta}) = 0] < 2^{-n}$$

*Note, breaking from convention, C will NOT have n input bits, but $\tilde{O}(|C|)$ input bits.*

**Proof.** Let $\alpha = \frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}$. We give a read once circuit $C_3$ which solves this. Note, the input length of $C_3$ will be greater than n and in fact will be roughly equal to the number of wires in $C_3$. When we later use $C_3$, the input bits will be randomly chosen from a smaller input giving us a super linear number of gates.

Let $C_1$ AND together $\ln_\epsilon(n^{-1-\alpha})$ bits. Then define $C_2$ to OR together $n^{1+\alpha}$ independent copies of $C_1$. Finally, let $C_3$ AND together $3n$ independent copies of $C_2$. Then $C_3$ circuits has $O(n^{2+\alpha})$ gates.

Now for $\delta < \epsilon$, we prove with exponentially low probability will $C_3$ fail to output 0 on $D_\delta$ or 1 on $D_{1-\delta}$. First observe that since $C_3$ is monotone, increasing the probability a bit is 1 will only increase the probability $C_3$ outputs 1. That is $\Pr[C_3(D_\delta) = 1] \leq \Pr[C_3(D_\epsilon) = 1]$. Similarly $\Pr[C_3(D_{1-\delta}) = 0] \leq \Pr[C_3(D_{1-\epsilon}) = 0]$. So it suffices to show it distinguishes $D_\epsilon$ from $D_{1-\epsilon}$.

First take an input from $D_\epsilon$. Then, the probability a random gate at each level will output 1 is:

- $C_1$ (AND):

$$\epsilon^{\ln_\epsilon(n^{-1-\alpha})} = \frac{1}{n^{1+\alpha}}$$

- $C_2$ (OR): See that for $k \geq 2$, $(1 - 1/k) \geq e^{-4/3k}$. Then:

$$1 - \left(1 - \frac{1}{n^{1+\alpha}}\right)^{n^{1+\alpha}} \leq 1 - e^{-4/3} \leq \frac{3}{4}$$

- $C_3$ (AND):

$$(3/4)^{3n} < 2^{-n}$$

Now take an input from $D_{1-\epsilon}$. For simplicity, we look at feeding $D_\epsilon$ into the negated circuit (ORs and ANDs swapped), which is cleaner. Then the probability of outputting a 0 on $C_3$ (or equivalently 1 on the negated circuit with negated input) is:

- $C_1$ (OR): Notice that by our choice of $\alpha$ that $\alpha = (1+\alpha)\frac{\ln(1-\epsilon)}{\ln(\epsilon)}$. Then:

$$1 - (1-\epsilon)^{\ln_\epsilon(n^{-1-\alpha})} = 1 - e^{\ln(1-\epsilon)\frac{-(1+\alpha)\ln(n)}{\ln(\epsilon)}} = 1 - n^{-\alpha} \leq e^{-n^{-\alpha}}$$

- $C_2$ (AND):

$$e^{-n^{-\alpha}n^{1+\alpha}} = e^{-n}$$

- $C_3$ (OR): For $n \geq 12$

$$1 - (1 - e^{-n})^{3n} \leq 3ne^{-n} < 2^{-n}$$

Thus with probability less than $2^{-n}$ will $C_3$ fail to distinguish $D_\epsilon$ from $D_{1-\epsilon}$ and the size of $C_3$ is order of the fan in of $C_2$ times the fan in of $C_3$, which is $O\left(n^{2+\alpha}\right) = O\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}}\right)$. ◀

Then feeding in random indexes from a $\mathrm{Maj}_\epsilon$ input and an averaging argument gives an $\epsilon$-promise majority circuit.

▶ **Theorem 26.** *For any $\epsilon \in (0, 1/2)$ there exists a monotone depth 3 AC0 circuit solving the $\epsilon$-promise majority problem with size $O\left(n^{2+\frac{\ln(1-\epsilon)}{\ln(\epsilon)-\ln(1-\epsilon)}}\right)$.*

**Proof.** Let $\alpha = \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}$. Take $C_3$ from Lemma 25. Construct a distribution of circuits C which is $C_3$ with inputs from $C_3$ mapped to random bits in our n bit input.

An input into C with $\epsilon n$ ones is equivalent to an input into $C_3$ from $D_\epsilon$. Thus for any $x \in \text{Maj}_\epsilon$, C outputs the wrong answer with probability less than $2^{-n}$. Thus by an averaging argument, some circuit in the distribution of circuits C always outputs the right answer. Then that circuit solves $\epsilon$-promise majority. ◀

### C.1.1   Ajtai's Circuit for Small Promises

We can improve our result for small $\epsilon$ by observing we only need to get failure probability below $|\text{Maj}_\epsilon|^{-1}$, which will be more than $2^{-n}$ for small $\epsilon$. This gives us:

▶ **Theorem 20.** *For any $\epsilon \in (0, 1/2)$ there exists a monotone depth 3 AC0 circuit solving the $\epsilon$-promise majority problem with size $O\left( (\epsilon \ln(\epsilon))^2 \, n^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}} \right)$.*

**Proof.** First, I show that

$$|\text{Maj}_\epsilon^0| < 2^{\log(2e/\epsilon)\epsilon n}$$

Consider sampling an element from $\text{Maj}_\epsilon^0$ by choosing $\epsilon n$ indexes out of $(1 + \epsilon)n$ bits. For any location chosen from the first n locations will be the locations of the 1 bits. See that this selects from every input in $\text{Maj}_\epsilon^0$ with some probability. Further see that this only samples at most $\binom{(1+\epsilon)n}{\epsilon n}$ values. Then we have:

$$|\text{Maj}_\epsilon^0| \leq \binom{(1+\epsilon)n}{\epsilon n} \leq \left( \frac{e(1+\epsilon)n}{\epsilon n} \right)^{\epsilon n} < \left( \frac{2e}{\epsilon} \right)^{\epsilon n} = 2^{\log(2e/\epsilon)\epsilon n}$$

Then we trivially get that $|\text{Maj}_\epsilon| \leq 2^{2\log(2e/\epsilon)\epsilon n}$.

Now using Lemma 25 with $n = 2\log(2e/\epsilon)\epsilon n$, we get a circuit C with size

$$O\left( (2e\epsilon \log(2e/\epsilon))^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}} \, n^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}} \right) = O\left( (\epsilon \ln(\epsilon))^2 \, n^{2 + \frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}} \right)$$

For convenience, I am assuming $\frac{\ln(1-\epsilon)}{\ln(\epsilon) - \ln(1-\epsilon)}$ is bounded. If not, then $\epsilon$ is close to one half, $\epsilon \ln(\epsilon)$ is constant, and the bound holds by Theorem 26.

So that for any $\delta < \epsilon$

$$\Pr[C(D_\delta) = 1] < 2^{-2\log(2e/\epsilon)\epsilon n} < |Maj_\epsilon|^{-1}$$

$$\Pr[C(D_{1-\delta}) = 0] < 2^{-2\log(2e/\epsilon)\epsilon n} < |Maj_\epsilon|^{-1}$$

Then applying C to random indexes of our input, any $x \in \text{Maj}_\epsilon$ will give the wrong answer with probability less than $|\text{Maj}_\epsilon|^{-1}$. Thus some choice of indexes works for every $c \in \text{Maj}_\epsilon$. Then C at these indexes solves $\epsilon$-promise majority. ◀

Given amplification, our depth 4 upper bounds in Theorem 5 follows in a similar manner.

These marginal improvements tell us that for polynomially small $\epsilon$, depth 3 $\epsilon$-promise majority circuits can be smaller than quadratic. This still leaves an $\epsilon$ factor between our lower bounds in Theorem 4 and our best known circuits in Theorem 20.

## C.2 Amplification

Now to get our smaller circuits, we need amplify a constant $\text{Maj}_\epsilon$ input to a $\text{Maj}_{1/\log(n)}$ input.

One way to perform this amplification is to use a randomized construction similar to Ajtai's that has a $1/\text{polylog}(n)$ chance of outputting correctly whether this is a $\text{Maj}_\epsilon^1$ input or a $\text{Maj}_\epsilon^0$ input. Then we can use a Chernoff bound to show that if we select $n\text{polylog}(n)$ of these circuits at random, with probability all but $2^{-n}$ a $\text{Maj}_\epsilon^1$ input will map to a $\text{Maj}_{1/\text{polylog}(n)}^1$ input and a $\text{Maj}_\epsilon^0$ input will map to a $\text{Maj}_{1/\text{polylog}(n)}^0$. Thus some circuit will do that mapping.

Instead, our amplification procedure will be a majority of short walks in an expander. This, combined with a Chernoff bound for expander graphs, will give us our amplification. Note this solution actually gives a significantly WORSE polylogarithmic overhead than the randomized construction above. But we are ignoring polylogarithmic factors in our result and we can use this same construction in our uniform bounds.

For notation:

▶ **Definition 27** (Random walks). *If $G$ is a graph, let $G_t$ be the length $t$ walks on $G$.*

The expander Chernoff bound is given explicitly by Healy [8] as theorem 1:

▶ **Lemma 28.** *For $G$ a regular graph with n vertices and spectral expansion $\lambda$ and $f : [n] \to [0,1]$ be any function. Let $\mu = \mathbb{E}_{v \in [n]}[f(v)]$. Then for any $\epsilon > 0$:*

$$\Pr_{w \in G_t}\left[\left|\mathbb{E}_{i \in [t]}[f(w_i)] - \mu\right| \geq \epsilon\right] \leq 2e^{-\frac{\epsilon^2(1-\lambda)t}{4}}$$

Then this almost directly gives our amplification .

▶ **Lemma 29.** *For constant $\epsilon \in (0, 1/2)$ and any $\delta \in (0, 1/2)$, there is a P-Uniform, depth 2 monotone AC0 circuit, C, outputting $O(n/\delta^{O(1)})$ bits such that:*

$$C\left(Maj_\epsilon^1\right) \subseteq Maj_\delta^1$$

$$C\left(Maj_\epsilon^0\right) \subseteq Maj_\delta^0$$

*And C has size*

$$|C| = O(n/\delta^{O(1)})$$

*Where we allow both constants to depend on $\epsilon$.*

**Proof.** Take a constant degree d regular expander graph G with constant spectral expansion $\lambda < 1$ over the n input bits. That such a graph exists is standard result [6]. Now we just need our walk to be long enough that we have only probability $\delta$ that our sample differs from the mean by more than $1/2 - \epsilon$. This happens when

$$t = \frac{4(\ln(2) - \ln(\delta))}{(1/2 - \epsilon)^2(1 - \lambda)} = \frac{16(\ln(2) - \ln(\delta))}{(1 - 2\epsilon)^2(1 - \lambda)}$$

For notation, let $M$ take an index to value of a bit at that index. See that for an input with at most $\epsilon$ fraction of the bits one, $\mathbb{E}_{v \in [n]}[M(v)] = \mu \leq \epsilon$ and:

$$\Pr_{w \in G_t}[\mathbb{E}_{i \in [t]}[M(w_i)] \geq 1/2] = \Pr_{w \in G_t}[\mathbb{E}_{i \in [t]}[M(w_i)] - \epsilon \geq 1/2 - \epsilon]$$

$$\leq \Pr_{w \in G_t}[\mathbb{E}_{i \in [t]}[M(w_i)] - \mu \geq 1/2 - \epsilon]$$

$$\leq \Pr_{w \in G_t}[|\mathbb{E}_{i \in [t]}[M(w_i)] - \mu| \geq 1/2 - \epsilon]$$

$$\leq 2e^{-\frac{(1/2-\epsilon)^2(1-\lambda)t}{4}}$$

$$= 2e^{\ln(\delta) - \ln(2)}$$

$$= \delta$$

Thus with probability at most $\delta$ will it give the wrong output. Thus taking all random walks on G of length t will give a $\text{Maj}_\delta^0$ input. Similarly for an input with at most $\epsilon$ 0s.

Thus such a random walk works, now we just need to show that there aren't too many walks. See that the number of random walks is:

$$|G_t| = n \cdot d^{t-1}$$

$$= n \cdot d^{\frac{4(\ln(2) - \ln(\delta))}{(1/2-\epsilon)^2(1-\lambda)} - 1}$$

$$= n \cdot d^{\frac{4\ln(2)}{(1/2-\epsilon)^2(1-\lambda)} - 1} \left(\frac{1}{\delta}\right)^{\frac{4\ln(d)}{(1/2-\epsilon)^2(1-\lambda)}}$$

$$= O(n) \left(\frac{1}{\delta}\right)^{O(1)}$$

Now for the size of the circuit computing majority for one of these walks. For each of these walks, we have a DNF computing the majority. Any DNF on t variables is of size at most

$$2^t = 2^{\frac{16(\ln(2)-\ln(\delta))}{(1-2\epsilon)^2(1-\lambda)}} = O(1)\,(1/\delta)^{O(1)}$$

Then the total size of the circuit is just that, times the number of walks, which is just $O(n/\delta^{O(1)})$. ◄

And the specific delta we need is inverse poly logarithmic. This immediately gives us Lemma 22.

▶ **Lemma 22.** *For any constant k, $\epsilon \in (0, 1/2)$, there exists a monotone, P-Uniform depth 3 circuit of size $\tilde{O}(n)$ amplifying a $\text{Maj}_\epsilon^0$ input to a $\text{Maj}_{\frac{1}{\ln(n)^k}}^0$ output and a $\text{Maj}_\epsilon^1$ input to a $\text{Maj}_{\frac{1}{\ln(n)^k}}^1$ output.*

## C.3 Recursive Majority Proof

We can get sub quadratic sized circuits for promise majority, down to size $n^{1+\alpha}$ for arbitrarily small $\alpha$ with a constant depth depending on $\alpha$. The basic idea is to first amplify the input. Then separate it into $n^\alpha$ sized sections and use our promise majority circuit on each of these. This will give us a new input which satisfies a slightly worse promise but is only of size $n^{1-\alpha}$. Then we can just repeat this process $\log_2(1/\alpha)$ times before the output is a single value which solves our promise majority problem.

This is the iterative equivalent of a recursive divide an conquer technique where we separate the input into $n^{\frac{1+\alpha}{2}}$ groups of size $n^{\frac{1-\alpha}{2}}$ and do our recursive promise majority on

each of these groups. Then do a final promise majority on the output. These produce the same circuits and the analysis is almost the same.

This process will work using any circuit computing promise majority as a subroutine as long as we can easily perform whatever amplification it needs. We use $1/\ln(n)$ for the $\epsilon$ used by our promise majority sub routine. This is convenient, but it could be smaller, even down to an appropriate polynomial as long as we can still do the amplification efficiently.

We first prove the strategy works on a $\mathrm{Maj}_{1/\mathrm{polylog}(n)}$ input, and this combined with amplification solves our promise majority problem.

▶ **Lemma 23.** *If there is a depth 3 alternating circuit with size $n^\alpha$ solving $1/\ln(n)$-promise majority, then for any positive integer $k$, there is a depth $1 + 2 \cdot k$ alternating circuit solving $1/\ln(n)^k$-promise majority with size:*

$$kn^{\frac{1}{1-\left(\frac{\alpha-1}{\alpha}\right)^k}}$$

*Which is uniform and monotone if the depth 3 circuit is uniform and monotone.*

**Proof.** The proof will go by induction. For the base case, $k = 1$, we have by assumption a size $n^\alpha = n^{\frac{1}{1-\frac{\alpha-1}{\alpha}}}$ depth 3 circuit.

In general, for $k \geq 2$, let $\beta = \frac{1}{1-\left(\frac{\alpha-1}{\alpha}\right)^k}$, and $\gamma = \frac{\beta-1}{\alpha-1}$. Separate the input into sets of size $n^\gamma$. Then let $C_1$ be the depth 3 circuit that just computes a $1/\ln(n)$-promise majority on each of these groups. See that the number of groups is $n^{1-\gamma}$. Then the size $C_1$ will be:

$$n^{1-\gamma}\left(n^\gamma\right)^\alpha = n^{1-\gamma+\alpha\gamma} = n^{1+(\alpha-1)\frac{\beta-1}{\alpha-1}} = n^\beta$$

By a counting argument, if an input satisfies the $1/\ln(n)^k$ promise, since our promise majority at depth 3 solves the $1/\ln(n)$ promise, the output of $C_1$ will satisfy a $1/\ln(n)^{k-1}$ promise. This is because fooling one of the promise majority will require $n^\gamma/\ln(n)$ of the minority bits. Since we only have $n/\ln(n)^k$ bits, we can only fool $n^{1-\gamma}/\ln(n)^{k-1}$ groups, which is only a $1/\ln(n)^{k-1}$ fraction of the groups.

Now we can use our induction hypothesis on this output. By assumption, we have a circuit $C_2$ with depth $1 + 2 \cdot (k-1)$ that solves the $1/\ln(n)^{k-1}$-promise majority on $n^{1-\gamma}$ bits and has size $(k-1)n^{\frac{1-\gamma}{1-\left(\frac{\alpha-1}{\alpha}\right)^{k-1}}}$. That is $C_2 \circ C_1$ solves the $1/\ln(n)^k$ promise majority problem. Now for the size of $C_2$. First we will simplify the number of groups. See that

$$
\begin{aligned}
1 - \gamma &= 1 - \frac{\beta-1}{\alpha-1} \\
&= \frac{\alpha-\beta}{\alpha-1} \\
&= \frac{\alpha - \frac{1}{1-\left(\frac{\alpha-1}{\alpha}\right)^k}}{\alpha-1} \\
&= \frac{\alpha\left(1-\left(\frac{\alpha-1}{\alpha}\right)^k\right)-1}{(\alpha-1)\left(1-\left(\frac{\alpha-1}{\alpha}\right)^k\right)} \\
&= \frac{\alpha-1-\alpha\left(\frac{\alpha-1}{\alpha}\right)^k}{(\alpha-1)\left(1-\left(\frac{\alpha-1}{\alpha}\right)^k\right)} \\
&= \frac{1-\left(\frac{\alpha-1}{\alpha}\right)^{k-1}}{1-\left(\frac{\alpha-1}{\alpha}\right)^k}
\end{aligned}
$$

Then the size is of $C_2$ is

$$(k-1)n^{\frac{1-\gamma}{1-\left(\frac{\alpha-1}{\alpha}\right)^{k-1}}} = (k-1)\left(n^{\frac{1-\left(\frac{\alpha-1}{\alpha}\right)^{k-1}}{1-\left(\frac{\alpha-1}{\alpha}\right)^{k}}}\right)^{\frac{1}{1-\left(\frac{\alpha-1}{\alpha}\right)^{k-1}}}$$

$$= (k-1)\left(n^{\frac{1}{1-\left(\frac{\alpha-1}{\alpha}\right)^{k}}}\right)$$

$$= (k-1)\left(n^{\beta}\right)$$

Then $C_2 \circ C_1$ has size

$$n^{\beta} + (k-1)n^{\beta} = kn^{\beta}$$

Further, see that for any promise input, if you flip all the bits, you get a promise input that outputs the opposite value. So switching all the AND gates to OR gates and all the OR gates to AND gates still gives a promise majority circuit. This can be seen by negating the circuit twice, once to use De Morgans, and once to flip all the input bits.

Thus, we can choose $C_1$ such that its top level of gates are the same as the bottom level gates of $C_2$, so they can merge. Thus $C_2 \circ C_1$ has depth $3 + 1 + 2 \cdot (k-1) - 1 = 1 + 2 \cdot k$. So $C_2 \circ C_1$ is our promise majority circuit as we wanted. ◀

## D  Viola's Promise Majority Circuit

Viola [16] gave a depth 3 circuit $\epsilon$-promise majority circuit where the lowest level has fan in $O\left(\frac{\ln(n)}{\ln(\ln(n))}\right)$ and solves the $1/\ln(n)$ promise majority problem. This combined with the amplification from Lemma 22 gives a constant depth circuit for $\epsilon$-promise majority.

Normally this would give a depth 4 circuit. But the second level of this circuit, only depends on $O(\ln(n))$ bits, so we can switch them from CNFs to DNFs or DNFs to CNFs with only polynomial overhead. This switching collapses a layer yielding a polynomial size depth 3 circuit.

Unfortunately, the depth 3 circuit is significantly larger than cubic, thus we will consider amplification separately. This leaves us is in much the same situation as the non uniform construction, except that our $\frac{1}{\ln(n)}$-promise promise majority circuit is larger, requiring greater depths to get the same size.

### D.1  Improvement on Viola's Promise Majority

First, we will give some intuition for Viola's promise majority. Viola's construction builds off of Lautemann's proof [11] that

$$\text{BPTime}(n) \subseteq \Sigma_2\text{Time}(n^2\text{polylog}(n))$$

We will look at a slight generalization of Lautemann's approach. Let $M : [n] \rightarrow \{0,1\}$ simply be the function that takes an index to the value of the input bit at that index. The idea is to take a family of bijections, $F$, and some collection of tuples of bijections $G \subseteq F^m$ and check if:

$$\exists f_1, ..., f_m \in G : \forall u \in [n] : \exists i \in [m] : M(f_i(u))$$

The idea is that each function spreads the input bits in some way, then we OR them together. Then we ask, was there some spreading that made our 1s cover the entire space? If we had less than a $1/m$ fraction of 1s, we can't do this. This is because at best our functions will multiply our number of ones by m.

If we have more than $1/m$ fraction of ones, sure SOME choice of functions will give us the spread we need. But we can't enumerate over ALL functions. Lautemann used bitwise xor with random strings as his family of functions and used a probabilistic argument to show this works. But every choice of m n bit strings is still too many options, so Viola used random walks on expander graphs over the indexes to get his values to xor with. We take this one step further and let the functions themselves be the random walks.

In the ones case, we argue that some choice of random walk will "spread" the ones to cover the space. The idea is to view the construction of the family of functions as randomized. Then we argue the probability over functions and indexes that we don't cover that index is less than the number of indexes. Thus, by an averaging argument, some family of functions actually obtains the average, and thus never outputs a 0.

That is, we want to prove

$$\Pr_{f_1,\ldots,f_m \in G, u \in [n]}[\forall i \in [m] : M(f_i(u)) = 0] < 1/n$$

Which implies:

$$\exists f_1, \ldots, f_m \in G : \forall u \in [n] : \exists i \in [m] : M(f_i(u))$$

▶ **Theorem 24.** *There exists a P-uniform, size $O\left(n^{3+o(1)}\right)$ depth 3 AC0 circuit solving the $\frac{1}{\ln(n)}$-promise majority problem.*

**Proof.** First, observe that there exists an expander, G, on n vertices (representing the n bits) that is $d = 5\ln(n)$ regular with spectral expansion $\lambda \leq \frac{2}{\sqrt{d}} = \frac{2}{\sqrt{5\ln(n)}}$. Specifically, we will use a bipartite Ramanujan graph as is shown to be constructable by Cohen [6]. To do this we will assume that n is even. Otherwise we can add a 0 bit and our promise will be changed negligibly. We will use the bipartite structure later for defining our random walks in such a way to make our functions bijections.

We want a way to enumerate the paths of length t over both the starting index, and the choice of steps in an independent way. Let $G_t$ be the length t walks over graph G. Let $W_d^t$ be the set of walk strategies for length t walks regardless of origin. That is the choices of edges to take on a length t path in a d regular graph. This will correspond to at every step, giving an index of the next edge to take. We will show later how to index the edges in such a way that each step forms a bijection.

Let $w \in W_d^t$ be a function $w : [n] \to G_t$ that takes a starting vertex and a walk strategy to that walk. Thus we can see that $G_t = \{w(i) : i \in [n], w \in W_d^t\}$, or that the set of all length t walks is equal to all walks from every starting vertex with every walk strategy.

We will let the length of our walks be $t = 2\ln(n)/\ln(\ln(n)) + 1$. Then we have number of walk strategies:

$$|W_d^t| = d^{t-1} = (5\ln(n))^{2\ln(n)/\ln(\ln(n))} = n^{2+o(1)}$$

As above, let $M$ be the function that takes a variable index to the value of that bit in the input. We claim that the proposition

$$\exists w \in W_d^t : \forall u \in [n] : \exists i \in [t] : M(w(u)_i)$$

is true if our input is in $\text{Maj}^1_{1/\ln(n)}$, and false if it is in $\text{Maj}^0_{1/\ln(n)}$. Finally, this statement is computed by a depth 3 AC0 circuit with size $O(n^{3+o(1)})$.

0   Input is in $\text{Maj}^0_{1/\ln(n)}$: This comes from a counting argument. But we need our functions, $w(u)_1, ..., w(u)_t$, to be permutations of the vertices, that is bijections. To do this, we make each choice of edge a permutation.

A walk strategy will just be a series of t numbers saying whether to take the first, second, up to dth edge. We label each edge with an index so that every vertex has exactly one edge with each label adjacent to it. Then to take edge i is just to go across the edge labelled i adjacent to the vertex. Then we need d disjoint perfect matchings on our graph.

Our graph is regular and bipartite, so it has a perfect matching. Take a perfect matching and label each edge in that matching 1 and remove them. This leaves us with again a bipartite regular graph, so we can get another matching for label 2, and so on until we have labeled every edge. By definition of perfect matching, each vertex and each label will have exactly one edge incident to that vertex with that label.

Then with this labeling, each step is a permutation, so every length $i$ walk strategy is a permutation. Therefore the number of pairs, $(u, i)$, such that $M(w(u)_i)$ is one is at most

$$\frac{n}{\ln(n)} \cdot t < \frac{n}{\ln(n)} \frac{3\ln(n)}{\ln(\ln(n))} < n$$

Therefore, for some u, there must not be any i so that $M(w(u)_i) = 1$. Thus there does not exist such a walk strategy so that for all u, that strategy starting at u hits a one.

1   Input is in $\text{Maj}^1_{1/\ln(n)}$: As a reminder, our strategy is to prove that the probability of outputting a 0 for every step of a random walk is so low that some strategy doesn't do this on any bit. That is:

$$\Pr_{w \in W^t_d, u \in [n]}[\forall i \in [m] : M(w(u)_i) = 0] < 1/n$$

This comes from the well known result [15] of expander walks that for any expander on n vertices $G$ with spectral expansion $\lambda$ and any set of vertices $A$ with density $\mu = |A|/n$:

$$\Pr_{w \in G_t}[\forall i \in [t] : w_i \in A] \le \mu(\mu + (1 - \mu)\lambda)^{t-1}$$

In particular, for our graph

$$\Pr_{w \in W^t_d, u \in [n]}[\forall i \in [m] : M(w(u)_i) = 0] = \Pr_{w \in G_t}[\forall i \in [m] : M(w_i) = 0]$$
$$\le \mu(\mu + (1 - \mu)\lambda)^{t-1}$$
$$< \left(1/\ln(n) + 2/\sqrt{5\ln(n)}\right)^{t-1}$$
$$< \left(1/\sqrt{\ln(n)}\right)^{t-1}$$
$$< \left(1/\sqrt{\ln(n)}\right)^{2\ln(n)/\ln(\ln(n))}$$
$$= n^{-1}$$

By an averaging argument, there must be some walk strategy w so that

$$Pr_{u \in [n]}[\forall i \in [m] : M(w(u)_i) = 0] < n^{-1}$$

But there are only n choices for u. Thus for all u, there is some $i$ so that $M(w(u)_i) = 1$. That is exactly what we needed to prove.

Finally, see the statement represents a small AC0 circuit. This is purely syntactic. Each quantifier is a layer of gates where the fan in is represented by the domain of the variables. 3 quantifiers means 3 layers, and the size is at most the product of number of choices at each level. Thus the size is at most:

$$|W_d^t| \cdot n \cdot t = n^{2+o(1)} \cdot n \cdot (2\ln(n)/\ln(\ln(n)) + 1) = n^{3+o(1)}$$

◄