

Nonuniform ACC Circuit Lower Bounds

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The class ACC consists of circuit families with constant depth over unbounded fan-in AND, OR, NOT, and MOD_m gates, where $m > 1$ is an arbitrary constant. We prove the following.

- NEXP, the class of languages accepted in nondeterministic exponential time, does not have nonuniform ACC circuits of polynomial size. The size lower bound can be slightly strengthened to quasipolynomials and other less natural functions.
- E^{NP} , the class of languages recognized in $2^{O(n)}$ time with an NP oracle, doesn't have nonuniform ACC circuits of $2^{n^{o(1)}}$ size. The lower bound gives an exponential size-depth tradeoff: for every d, m there is a $\delta > 0$ such that E^{NP} doesn't have depth- d ACC circuits of size 2^{n^δ} with MOD_m gates.

Previously, it was not known whether EXP^{NP} had depth-3 polynomial-size circuits made out of only MOD_6 gates. The high-level strategy is to design faster algorithms for the circuit satisfiability problem over ACC circuits, then prove that such algorithms entail these lower bounds. The algorithms combine known properties of ACC with fast rectangular matrix multiplication and dynamic programming, while the second step requires a strengthening of the author's prior work.

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1. INTRODUCTION

Nonuniform computation allows the sizes of programs to grow with the sizes of inputs. A nonuniform computation can be naturally represented as an infinite family of *Boolean circuits*, one for each possible input length. A long-term goal of complexity theory is to understand how nonuniform computations compare with the usual uniform models, which have fixed-size programs. Nonuniform circuit families can recognize *arbitrary* (even undecidable) languages, by having a large enough circuit for each input length. Nonuniformity becomes much more interesting when we restrict the *sizes* of computations: are there interesting uniform computations that cannot be simulated by

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small nonuniform circuit families? For instance, could every *exponential-time* uniform computation be simulated using circuit families of only polynomial size? Amazingly, this question is still open. For another example, $P \neq NP$ follows if one could provide an NP problem that cannot be solved by any circuit family where the size of the n th circuit is at most polynomial in n . Lower bounds against nonuniform computation establish impossibility results for computation in the physical world: it could be that $P \neq NP$, yet NP-complete problems can still be efficiently solved using “bloated” programs with sufficiently many lines of code for large inputs. Nonuniform circuit-size lower bounds for NP would rule out this possibility. (However, the best-known circuit-size lower bound for any NP problem is only $5n - o(n)$ [Iwama and Morizumi 2002; Lachish and Raz 2001].)

In the early 1980’s, researchers began to carefully study the power of nonuniform *low-depth* circuits. Intuitively, such circuits correspond to extremely fast parallel computations. The initial hope was that if some functions in NP were proved to require large, restricted circuit families, then by gradually lifting the restrictions over time, superpolynomial-size unrestricted lower bounds for NP could be attained, proving $P \neq NP$. Furst et al. [1984], and independently Ajtai [1983], showed that functions such as the parity of n bits cannot be computed by polynomial-size AC^0 circuits, that is, polynomial size circuit families of constant depth over the usual basis of AND, OR, and NOT gates, where each AND and OR may have arbitrarily many inputs. Yao [1985] improved the lower bounds to exponential size, and Håstad [1986] proved essentially optimal AC^0 lower bounds for parity. Around the same time, Razborov [1985] proved superpolynomial lower bounds for solving clique with *monotone* circuits (i.e., general circuits without NOT gates), and the bound was improved to exponential size by Alon and Boppana [1987]. However, it was later shown [Razborov 1989] that these techniques probably would not extend to general circuits.

Encouraged by the progress on AC^0 , attention turned to lower bounds for what seemed to be minor generalizations. The most natural generalization was to grant AC^0 the parity function for free. Razborov [1987] proved an exponential lower bound for computing the majority of n bits with constant-depth circuits made up of AND, OR, NOT, and MOD_2 gates. (A MOD_m gate outputs 1 iff m divides the sum of its inputs.) Then, Smolensky [1987] proved exponential lower bounds for computing MOD_q with constant-depth circuits made up of AND, OR, NOT, and MOD_p gates, for distinct primes p and q . Barrington [1989] suggested the next step would be to prove lower bounds for the class ACC, which consists of constant-depth circuit families over the basis AND, OR, NOT, and MOD_m for arbitrary constant $m > 1$.¹ It is here that progress on strong circuit lower bounds began to falter (although there has been progress on further restricted cases, cf. the Preliminaries). Although it was conjectured that the majority of n bits cannot have polynomial-size ACC circuits, strong ACC lower bounds remained elusive.

After some years of failing to prove a superpolynomial lower bound, the primary questions were weakened. Rather than trying to find simple functions that cannot be computed with weak circuits, perhaps we could rule out weak circuits for complicated functions. Could one prove that nondeterministic *exponential* time (NEXP) doesn’t have polynomial size circuits? A series of papers starting with Nisan and Wigderson [Babai et al. 1993; Impagliazzo et al. 2002; Klivans and van Melkebeek 2002; Nisan and Wigderson 1994] showed that even this sort of lower bound would imply derandomization results: in the case of NEXP lower bounds, it would imply that Merlin-Arthur

¹The class is also called ACC^0 in the literature. However, as ACC^i is hardly studied at all, for any $i > 0$, at the present time it makes sense to drop the superscript.

games can be nontrivially simulated with nondeterministic algorithms. This indicated that proving good circuit lower bounds for NEXP would already require significantly new ideas.

In this article, we address two frontier questions concerning nonuniform circuit complexity.

- (1) Does nondeterministic $2^{O(n)}$ time have nonuniform polynomial size ACC circuits? (Is $\text{NTIME}[2^{O(n)}]$ in nonuniform ACC?)
- (2) Does exponential time with an NP oracle have nonuniform polynomial-size circuits? (Is $\text{EXP}^{\text{NP}} \subseteq \text{P/poly}$?)

Over the years, these questions have turned into notorious and somewhat embarrassing open problems, because it seems so obvious that the answers should be no. It was open if EXP^{NP} could be recognized with depth-3 polynomial size circuits made out of only MOD_6 gates.² We make headway on these frontiers, giving a strong *no* answer to the first question.

THEOREM 1.1. *$\text{NTIME}[2^n]$ does not have nonuniform ACC circuits of polynomial size.*

Stronger-size lower bounds hold (e.g., quasipolynomial size) but the results are not very clean; see Section 5.1 for details. For EXP^{NP} , we can prove exponential lower bounds.

THEOREM 1.2 (EXPONENTIAL SIZE-DEPTH TRADEOFF). *For every d and m , there is a $\delta > 0$ and a language in E^{NP} that fails to have nonuniform ACC circuits of depth d and size 2^{n^δ} with MOD_m gates.*

Recall that the smallest complexity class for which we know exponential-size (unrestricted) circuit lower bounds is $\text{EXP}^{\text{NP}^{\text{NP}}}$, the third level of the exponential hierarchy [Miltersen et al. 1999].

Extending the approach of this article to settle the second frontier question may be difficult, but this prospect does not look as implausible as it did before. If polynomial unrestricted circuits could be simulated by subexponential ACC circuits, or if one could improve just a little on the running time of algorithms for the circuit satisfiability problem, the second question would be settled.

1.1. An Overview of the Proofs

Let us sketch how these new lower bounds are proved, giving a roadmap for the rest of the paper. In recent work [Williams 2013], the author suggested a research program for proving nonuniform circuit lower bounds for NEXP. It was shown that for many circuit classes \mathcal{C} , sufficiently faster satisfiability algorithms for \mathcal{C} -circuits would entail nonuniform lower bounds for \mathcal{C} -circuits. The objective of this article is to carry out the proposed research program in the case of ACC circuits.

The proof of the lower bound for E^{NP} (Theorem 1.2) is a combination of complexity-theoretic ideas (time hierarchies, compression by circuits, the local checkability of computation) and algorithmic ideas (fast matrix multiplication, dynamic programming, table lookup).

First, we show that satisfiability algorithms for subexponential size n -input ACC circuits with running time $O(2^n/n^k)$ imply exponential size ACC lower bounds for E^{NP}

²Note that slightly larger classes such as MAEXP and NEXP^{NP} are known to not have polynomial-size circuits; see the Preliminaries.

(Theorem 3.2), where k is sufficiently large. (The model of computation for the satisfiability algorithm is flexible; we may assume the multitape Turing machine or a random access machine. See the Preliminaries.) This step considerably strengthens earlier work, which could only show that an $o(2^{n/3})$ -time algorithm for ACC circuit satisfiability implies lower bounds [Williams 2013]. The idea is to prove that, if there is a faster algorithm for ACC Circuit SAT, and there are subexponential ($2^{n^{o(1)}}$) size ACC circuits for E^{NP} , then every $L \in NTIME[2^n]$ can be accepted by a nondeterministic algorithm in $O(2^n n^{10}/n^k)$ time. (Here, 10 is a substitute for a small universal constant.) When $k > 10$, this contradicts the nondeterministic time hierarchy theorem [Seiferas et al. 1978; Žák 1983], so one of the assumptions must be false.

Two known facts are applied in the proof. First, there is a polynomial-time reduction from any $L \in NTIME[2^n]$ to the NEXP-complete problem SUCCINCT 3SAT such that every instance x of length n (for sufficiently large n) is reduced to a (unrestricted, not ACC) circuit C_x of size $O(n^5)$ with at most $n + 5 \log n$ inputs (Fact 1). That is, the string obtained by evaluating C_x on its $O(2^n n^5)$ possible assignments (in lex order) encodes a 3CNF formula F_{C_x} that is satisfiable iff $x \in L$. Informally, this says that languages $L \in NTIME[2^n]$ have “succinct” reductions to exponentially long 3SAT instances.

Second, if E^{NP} is in subexponential-size ACC, then (given an x) there is some satisfying assignment to the formula encoded by C_x that can be described by an ACC circuit W of subexponential size (Fact 2). That is, the string obtained by evaluating W on all possible assignments encodes a satisfying assignment to the exponentially long F_{C_x} . Informally, this means that, if E^{NP} has subexponential ACC circuits, then every “succinct” satisfiable 3SAT instance has at least one “succinct” satisfying assignment: compressible satisfiable formulas have compressible satisfying assignments.

We can try to combine these two facts, as follows. If C_x were an ACC circuit, then any $L \in NTIME[2^n]$ could be accepted in $O(2^n n^5/n^k)$ nondeterministic time, by guessing a subexponential ACC circuit W and constructing an ACC circuit satisfiability instance D built of C_x and W , where D is satisfiable if and only if W does not encode a satisfying assignment to F_{C_x} (as shown in the author’s prior paper [Williams 2013]). The circuit D has at most $n + 5 \log n$ inputs and $2^{n^{o(1)}}$ size, so the assumed ACC satisfiability algorithm can handle D in $O(2^n n^5/n^k)$ time.

The aforementioned argument does not quite work, because we do not know how to produce a C_x that is an ACC circuit (indeed, it may not be possible). An ACC SAT algorithm will not work on D , because D contains a copy of an unrestricted circuit C_x . However, assuming P has subexponential ACC circuits, we show how to *guess and verify* an equivalent ACC circuit C'_x in nondeterministic $O(2^n n^{10}/n^k)$ time using a slightly faster ACC SAT algorithm (Lemma 3.4). This makes it possible to prove ACC lower bounds even with weak ACC satisfiability algorithms. Furthermore, this part of the proof does not use specific properties of ACC, so it may be useful for proving stronger lower bounds in the future.

Second, we show how satisfiability of subexponential ACC circuits of depth d and n inputs can be determined in $2^{n-\Omega(n^\delta)}$ time, for a $\delta > 0$ that depends on d (Theorem 4.5). Given any such circuit C , replace it with C' which is an OR of 2^{n^δ} copies of C , where the first n^δ inputs of each copy are substituted with a variable assignment. This ACC circuit C' has $n - n^\delta$ inputs, $2^{O(n^\delta)}$ size, and C is satisfiable if and only if C' is. Applying a powerful result of Yao, Beigel-Tarui, and Allender-Gore (Lemma 4.1), C' can be replaced by an equivalent depth-2 circuit C'' of $2^{n^{\delta 2^{O(d)}}}$ size, which consists of an efficiently computable symmetric function at the output gate and AND gates below

it. Setting $\delta \ll 1/2^{O(d)}$, and exploiting the structure of the depth-2 circuit, C'' can be evaluated on all of its possible assignments in $2^{n-n^\delta} \text{poly}(n)$ time (Lemma 4.2). This concludes the sketch of the E^{NP} lower bound.

The only use of the full assumption “ E^{NP} has ACC circuits” is in Fact 2. The lower bound for NEXP (Theorem 1.1) applies the result (which follows from work of Impagliazzo et al. [2002]) that if NEXP has polynomial size (unrestricted) circuits then satisfiable instances of SUCCINCT 3SAT already have polynomial size (unrestricted) circuits W encoding satisfying assignments (Theorem 5.2). But if P has ACC circuits, it is easy to see that these unrestricted circuits must have equivalent ACC circuits as well (Lemma 5.4). This helps extend the E^{NP} lower bound to NEXP. However, the resulting size lower bound is not exponential: from $S(n)$ -size circuits for NEXP, one only obtains $S(S(S(n)^c)^c)^c$ -size ACC circuits encoding satisfying assignments. This allows for some “half-exponential” type improvements in the size lower bounds against NEXP.

Perhaps the most interesting aspect of the proofs is that only the satisfiability algorithm for ACC circuits relies on specific properties of ACC. Improved exponential-time algorithms for satisfiability are the only barrier to further progress on circuit lower bounds for NEXP. In general, this article weakens the algorithmic assumptions necessary to prove lower bounds, and strengthens the lower bounds obtained. Let \mathcal{C} be a class of circuit families that is closed under composition (the composition of two circuit families from \mathcal{C} is also a family in \mathcal{C}) and contains AC^0 . Possible \mathcal{C} include constant-depth threshold circuits, Boolean formulas, and unrestricted Boolean circuits. The arguments of Section 3 and Section 5 imply the following metatheorem.

THEOREM 1.3. *There is a $k > 0$ such that, if satisfiability of \mathcal{C} -circuits with n variables and n^c size can be solved in $O(2^n/n^k)$ time for every c , then $\text{NTIME}[2^n]$ doesn't have nonuniform polysize \mathcal{C} -circuits.*

2. PRELIMINARIES

We presume the reader has background in circuit complexity and complexity theory in general. The textbook of Arora and Barak [2009] covers all the necessary material; in particular, Chapter 14 gives an excellent summary of ACC and the frontiers in circuit complexity.

On the Machine Model. An important point about this article is that the choice of uniform machine model is not crucial to the arguments. We show that if large classes have small nonuniform ACC circuits, then $\text{NTIME}[2^n] \subseteq \text{NTIME}[o(2^n)]$ (in fact, $\text{NTIME}[2^n] \subseteq \text{NTIME}[o(2^n/n^k)]$ for sufficiently large k), which is a contradiction in all computational models we are aware of. Moreover, Gurevich and Shelah [1989] proved that the nondeterministic machine models are tightly related in their time complexities. For example, let $\text{NTIME}_{\text{RTM}}[t(n)]$ be the languages recognized by nondeterministic $t(n)$ time random-access Turing machines, and let $\text{NTIME}_{\text{TM}}[t(n)]$ be the class for multitape Turing machines.

THEOREM 2.1 (GUREVICH AND SHELAH [1989]).

$$\bigcup_{c>0} \text{NTIME}_{\text{RTM}}[n \log^c n] = \bigcup_{c>0} \text{NTIME}_{\text{TM}}[n \log^c n].$$

As a consequence, even if we showed $\text{NTIME}_{\text{TM}}[2^n] \subseteq \text{NTIME}_{\text{RTM}}[2^n/n^k]$ for sufficiently large k , we would still obtain the desired contradiction. (Note that such a result

is not known for the deterministic setting.) A random access Turing machine can also simulate a standard random access machine with only constant factor overhead [Paul and Reischuk 1980]. Hence, in our proof by contradiction, we may assume that the source algorithm we're simulating is only a multitape TM, while the target algorithm has all the power we need to perform typical computations from the literature.

Notation. Inside of an algorithm description, the integer n refers to the length of the input to the algorithm. For a function $f : \mathbb{N} \rightarrow \mathbb{N}$, we use $\text{poly}(f(n))$ to denote a growth rate of the form $cf(n)^c$ for a constant c .

The size of a circuit refers to the number of wires in it. However, since attention shall be restricted to circuits with at least polynomially many gates, the distinction between the number of wires and gates does not matter. In this article, an *unrestricted circuit* has gate types AND/OR/NOT, unrestricted depth, and each gate has fan-in two. (That is, an unrestricted circuit is the generic variety used in the definition of P/poly .) All circuit size functions S considered in this article are assumed to be monotone nondecreasing, that is, $S(n+1) \geq S(n)$ for all n .

We say that a *circuit class* \mathcal{C} is a collection of circuit families that (a) *contains* AC^0 (for every circuit family in AC^0 , there is an equivalent circuit family in \mathcal{C}) and (b) *is closed under composition*: if $\{C_n\}$ and $\{D_n\}$ are families in \mathcal{C} , then for every c , the circuit family consisting of circuits which take n bits of input, feed them to $n^c + c$ copies of circuits from C_n , and feed the outputs to the inputs of D_{n^c+c} , is also a circuit family in \mathcal{C} . Essentially, all classes studied extensively in the literature (AC^0 , ACC , TC^0 , NC^1 , NC^2 , P/poly , etc.) are circuit classes in this sense. For classes that allow for superpolynomial-size circuits, the polynomial “ $n^c + c$ ” may be relaxed appropriately.

For a complexity class \mathcal{C} , the class i.o.- \mathcal{C} consists of languages $L \subseteq \Sigma^*$ such that there is a language $L' \in \mathcal{C}$ where $L \cap \Sigma^n = L' \cap \Sigma^n$ holds for infinitely many n .

When the expression “ $O(1)$ ” appears inside of the time bound for a complexity class, this is shorthand for the union of all classes where the $O(1)$ is substituted by a fixed constant. For example, the class $\text{TIME}[2^{n^{O(1)}}]$ is shorthand for $\bigcup_{c \geq 0} \text{TIME}[2^{n^c}]$.

Other Prior Work. Kannan [1982] showed that for any superpolynomial constructible function $S : \mathbb{N} \rightarrow \mathbb{N}$, the class $\text{NTIME}[S(n)]^{\text{NP}}$ does not have polynomial size circuits. Another somewhat small class known to not have unrestricted polynomial size circuits is MAEXP [Buhrman et al. 1998]. Later, it was shown that the MAEXP lower bound can be improved to half-exponential size functions f which satisfy $f(f(n)) \geq 2^n$ [Miltersen et al. 1999]. Kabanets and Impagliazzo [2004] proved that NEXP^{RP} either doesn't have polynomial-size Boolean circuits (over AND, OR, NOT), or it doesn't have polynomial-size arithmetic circuits (over the integers, with addition and multiplication gates). Note that $\text{NEXP}^{\text{RP}} \subseteq \text{MAEXP}$.

A line of work stemming from Toda's Theorem [Toda 1991] has studied ways of representing low-depth ACC circuits by certain depth-two circuits which will play a critical role in this article. Define a SYM^+ circuit to be a depth-two circuit which computes some symmetric function at the output gate, and computes ANDs of input variables on the second layer.³ Extending work on AC^0 by Allender [1989], Yao [1990] showed that every ACC circuit of s size can be represented by a probabilistic SYM^+ circuit of $s^{O(\log^c s)}$ size, where c depends on the depth, and the ANDs have $\text{poly}(\log s)$ fan-in. Beigel and Tarui [1994] showed how to remove the probabilistic condition. Allender

³Some have defined SYM^+ circuits differently, allowing the ANDs to take negated variables as input. In Appendix B, we sketch why there is no essential difference between the two definitions (for our purposes, at least).

and Gore [1991] showed that every subexponential *uniform* ACC circuit family can be simulated by subexponential *uniform* SYM^+ circuits. This was applied to show that the Permanent does not have *uniform* ACC circuits of subexponential size. Later, Allender [1999] improved the Permanent lower bound to polynomial size uniform TC^0 circuits. However, these proofs require uniformity, and the difference between uniformity and non-uniformity may well be vast (e.g., it is clear that $\text{P} \neq \text{NEXP}$, but open whether $\text{NEXP} \subseteq \text{P/poly}$). Green et al. [1995] showed that the symmetric function can be assumed to be the specific function which returns the *middle bit* of the sum of its inputs. This representation may also be used in the lower bounds of this article.

There has also been substantial work on representing ACC in other interesting ways [Agrawal et al. 2000; Barrington and Thérien 1988; Hansen 2006; Hansen and Koucký 2010] as well as many lower bounds in restricted cases [Barrington and Straubing 1995; Barrington et al. 1990; Caussinus 1996; Chattopadhyay and Wigderson 2009; Chattopadhyay et al. 2006; Grolmusz 1998; Grolmusz and Tardos 2000; Krause and Pudlák 1997; Thérien 1994; Yan and Parberry 1994]. Significant work has gone into understanding the *constant degree hypothesis* [Barrington et al. 1990] that a certain type of low-depth ACC circuit requires exponential size to compute the AND function. The hypothesis is still open.

All prior works on nonuniform ACC lower bounds attack the problem in a “bottom-up” way. (The exceptions are the uniform results mentioned previously [Allender 1999; Allender and Gore 1991].) Lower bounds have been proved for highly restricted circuits and these restrictions have been very gradually relaxed over time. In this article, the strategy is “top-down”: the goal is to find the smallest complexity classes for which it is still possible to prove superpolynomial ACC lower bounds. This is in line with the overall goal of eventually proving large circuit lower bounds for NP.

As mentioned before, this article builds on the author’s prior work, which showed that mild improvements over exhaustive search can sometimes imply lower bounds. Let us briefly review the prior state-of-the-art for Circuit SAT algorithms. It is known that CNF satisfiability can be solved in $2^{n-\Omega(n/\ln(m/n))} \text{poly}(m)$ time, where m is the number of clauses and n is the number of variables [Calabro et al. 2006; Dantsin and Hirsch 2009; Schuler 2005]. Recent work of Calabro-Impagliazzo-Paturi [Calabro et al. 2009] and Impagliazzo-Matthews-Paturi [Impagliazzo et al. 2012] has culminated in a randomized algorithm for depth- d AC^0 circuit satisfiability that runs in $2^{n-\Omega(n/(\log s)^{d-1})}$ time on circuits with s gates. Santhanam [2010] has applied ideas inspired by formula size lower bounds to show that for a fixed constant k , satisfiability of formulas over AND, OR, and NOT can be determined in $O(2^{n-n/c^k})$ time on n -variable formulas of size cn . Very recently, Seto and Tamaki [2013] have extended Santhanam’s results to formulas over arbitrary two-bit Boolean-valued functions. Unfortunately, these upper bounds are not yet strong enough to derive new circuit lower bounds for NEXP.

How Do We Avoid the Barriers? There are several well-known formal barriers to proving lower bounds. Let us say a little about their relation to this work. Intuitively, we circumvent the natural proofs barrier [Razborov and Rudich 1997] because of the use of diagonalization. More precisely, we rely heavily on strong completeness properties of a specific NEXP language, namely SUCCINCT 3SAT, to prove that it cannot have small ACC circuits. So it looks unlikely that one may extract any P-natural or NP-natural properties from the proof. (Furthermore, there is little evidence that ACC contains pseudorandom functions, so natural proofs may not be a barrier for ACC after all.) It is hard to formally rule out that a proof cannot possibly be made natural, without showing either an algorithmic lower bound (there is no efficient algorithm with certain properties) or a circuit upper bound (the circuit class under discussion has

pseudorandom functions). Nevertheless, statements like “Satisfiability of \mathcal{C} circuits is in $O(2^n/n^{10})$ time” do appear to be far weaker than statements like “There are no strong pseudorandom functions implementable with \mathcal{C} circuits”.

More conclusively, the approach of this work definitely avoids relativization [Baker et al. 1975] and algebrization [Aaronson and Wigderson 2009] because there are oracles A relative to which $\text{NEXP}^A \subset \text{ACC}^A$, and even $\text{NEXP}^{\bar{A}} \subset \text{ACC}^A$ (Scott Aaronson, personal communication). Here the lower bounds rely on a more efficient ACC SAT algorithm which uses nonrelativizing properties of ACC circuits. In general, the approach of using slightly faster SAT algorithms to prove lower bounds appears fruitful for circumventing oracle-based barriers, because all known improved SAT algorithms break down when oracles (or algebraic extensions thereof) are added to the instance. That is, significant improvements over exhaustive search necessarily exploit structure in instances that black-box methods cannot see.

3. A STRENGTHENED CONNECTION BETWEEN SAT ALGORITHMS AND LOWER BOUNDS

In this section, we prove that if one can achieve a very minor improvement over exhaustive search in satisfying ACC circuits, then one can prove lower bounds for ACC. The required improvement is so minor that we are able to achieve it, in the next section. However, let us stress upfront that all the results in this section hold equally well for other circuit classes as well: we only require basic properties of ACC that practically all robust circuit classes satisfy.

Define the ACC CIRCUIT SAT problem to be: given an ACC circuit C , is there an assignment of its inputs that makes C evaluate to 1? In recent prior work [Williams 2013], the author proved a relation between algorithms for ACC CIRCUIT SAT and lower bounds for ACC circuits.⁴

THEOREM 3.1 [WILLIAMS 2013]. *Let $s(n) = \omega(n^k)$ for every k . If ACC CIRCUIT SAT instances with n variables and n^k size can be solved in $O(2^{n/3}/s(n))$ time for every k , then E^{NP} does not have nonuniform ACC circuits of polynomial size.*

We shall sharpen this theorem considerably. Throughout the following, let $S : \mathbb{N} \rightarrow \mathbb{N}$ and $d : \mathbb{N} \rightarrow \mathbb{N}$ be monotone nondecreasing functions such that $S(n) \geq n$ and $d(n) \geq 1$. Let \mathcal{C} be a circuit class as defined in the Preliminaries. (\mathcal{C} can be ACC, TC^0 , NC^1 , P/poly , etc.) Define the \mathcal{C} -CIRCUIT SAT problem to be: Given a circuit C from class \mathcal{C} , is there an assignment of its inputs that makes C evaluate to 1?

THEOREM 3.2. *Let $S(n) \leq 2^{n/4}$. There is a $c > 0$ and $b > 0$ such that, if \mathcal{C} -CIRCUIT SAT instances with at most $n + c \log n$ variables, depth at most $2 \cdot d(3n) + b$, and $O(n \cdot S(2n) + S(3n))$ size can be solved in $O(2^n/n^c)$ time, then E^{NP} does not have nonuniform \mathcal{C} circuits of depth $d(n)$ and $S(n)$ size.*

The constant k depends on the model of computation in which the SAT algorithm is implemented, but for all typical models, c is not large (less than 10). For us, the important corollary is this: if ACC satisfiability has a slightly faster algorithm on circuits that are mildly larger than $S(n)$, then E^{NP} does not have ACC circuits of $S(n)$ size. In what follows, we prove Theorem 3.2 only for ACC circuits, but the proof also works for any other circuit class. (The reader can verify that the only two properties of ACC used are that the class contains AC^0 , and the class is closed under composition of circuit families.)

⁴In fact, a more general result for any circuit class was proved, which implies Theorem 3.1.

To understand the difficulty behind proving Theorem 3.2, let us recall the proof of Theorem 3.1 to see why it needed such a strong assumption. The generic proof idea in Williams [2013] for results such as Theorem 3.1 is to derive a contradiction from assuming small circuits for E^{NP} and a faster algorithm for CIRCUIT SAT. In particular, it is shown that under the two assumptions, every language $L \in NTIME[2^n]$ can be recognized in $NTIME[o(2^n)]$, which is false by the nondeterministic time hierarchy theorem [Seiferas et al. 1978; Žák 1983]. The contradiction is derived from stitching together several facts about circuits and satisfiability.

Define SUCCINCT 3SAT as the problem: Given a circuit C on n inputs, let F_C be the 2^n -bit instance of 3-SAT obtained by evaluating C on all of its possible inputs in lexicographical order. Is F_C satisfiable?

That is, given a *compressed encoding* of a 3-CNF formula, the task is to determine if the underlying decompressed formula is satisfiable. For natural reasons, call F_C the *decompression* of C , and call C the *compression* of F_C . The SUCCINCT 3SAT problem is a canonical NEXP-complete problem [Papadimitriou and Yannakakis 1986].

FACT 1. *There is a constant $c > 0$ such that for every $L \in NTIME[2^n]$, there is a reduction from L to SUCCINCT 3SAT which on input x of length n runs in $\text{poly}(n)$ time and produces a circuit C_x with at most $n + c \log n$ inputs and $c \cdot n^c + c$ size, such that $x \in L$ if and only if the decompressed formula F_{C_x} of $2^n \cdot \text{poly}(n)$ size is satisfiable.*

Fact 1 follows from several prior works concerned with the complexity of the Cook-Levin theorem.

THEOREM 3.3 [FORTNOW ET AL. 2005; TOURLAKIS 2001]. *There is a $c > 0$ such that for all $L \in NTIME[n]$, L reduces to 3SAT in $c \cdot n(\log n)^c$ time. Moreover, there is an algorithm A_L (with random access to its input) that, given an instance of L with length n and an integer $i \in [cn(\log n)^c + c]$ in binary, A_L outputs the i th clause of the resulting 3SAT formula in $c(\log n)^c + c$ time.*

In fact, the proofs in this references build on even earlier work [Cook 1988; Gurevich and Shelah 1989; Robson 1991; Schnorr 1978]. In a nutshell, all of these proofs exploit the *locality of computation*: every nondeterministic computation running in linear time can be represented with a nondeterministic circuit of size $O(n \cdot \text{poly}(\log n))$ which has a highly regular and efficiently computable structure. This circuit can be easily modeled as a 3-CNF formula using the Tseitin transformation [Tseitin 1968] that assigns a variable to each circuit wire and uses 3-CNF clauses to model the input-output relationships for each gate.

The value of c in Theorem 3.3 depends on the underlying computational model; typically one can take c to be at most 4. A standard padding argument (substituting 2^n in place of n) yields Fact 1. In more detail, given $L \in NTIME[2^n]$, we apply Theorem 3.3 to the language $L' = \{x01^{2^{|x|}} \mid x \in L\}$, which is in $NTIME[n]$. On an input x , this generates an equivalent 3SAT instance of length $O(2^{|x|}|x|^c)$. As it is easy to simulate random accesses to an input of the form $x01^{2^{|x|}}$ with a uniform $\text{poly}(|x|)$ size circuit, one can simulate the $O((\log n)^c)$ time algorithm of Theorem 3.3 on L' , with a uniform $\text{poly}(|x|^c)$ size circuit.

Using Fact 1, one can then prove that if the lower bound for E^{NP} fails to hold, then every succinctly compressible satisfiable formula that is output by the SUCCINCT 3SAT reduction has some succinctly compressible *satisfying assignment*.

FACT 2. *If E^{NP} has ACC circuits of size $S(n)$, then there is a fixed constant c such that for every language $L \in NTIME[2^n]$ and every $x \in L$ of length n , there is an ACC circuit W_x of size at most $S(3n)$ with $k \leq n + c \log n$ inputs such that the variable*

assignment $z_i = W(i)$ for all $i = 1, \dots, 2^k$ is a satisfying assignment for the formula F_{C_x} , where C_x is the circuit obtained by the reduction in Fact 1.

PROOF OF FACT 2. Consider the E^{NP} machine.

$N(x, i)$: Compute the SUCCINCT 3SAT reduction from x to C_x in polynomial time. Decompress C_x , obtaining a formula F of $O(2^{|x|}|x|^c)$ size. Let k be the number of inputs to C_x . Binary search for the lexicographically smallest satisfying assignment A to F , by repeatedly querying: *given (F, A) where $|A| \leq 2^k$, is there an assignment $A' \leq A$ that satisfies F ?* Then output the i th bit of A .

Note the queries can be answered in NP, and N needs $O(2^k)$ queries to the oracle. By assumption, N has ACC circuits of size $S(n)$. It follows that, for every $x \in L$, there is some satisfying assignment to F which is encoded by a circuit of size $S(|\langle x, i \rangle|) \leq S(3|x|)$, where $\langle \cdot, \cdot \rangle$ is a polynomial-time computable pairing function. \square

With these two facts, we may try to recognize any $L \in \text{NTIME}[2^n]$ with a $o(2^n)$ nondeterministic algorithm (a contradiction), as follows. Given a string x of length n , compute the SUCCINCT 3SAT circuit C_x in polynomial time and nondeterministically guess a $S(3n)$ -size circuit W . Now the goal is to check that W succinctly encodes a satisfying assignment for the underlying formula F_{C_x} . To verify this condition, the algorithm constructs a CIRCUIT SAT instance D . The circuit D has $n + c \log n$ inputs fed to $O(n)$ copies of C_x , so that when i is input to D , the copies altogether print the i th clause of the 3CNF formula F_{C_x} . These copies output three variable indices of length at most $n + c \log n$, along with sign bits (whether or not the variables are negated in the clause). Then, D feeds each index to a copy of W , which prints a bit. Finally, D compares the sign bits with the three bits printed by the copies of W , and outputs 0 iff the variable assignment encoded by W satisfies the i th clause. Observe D has $\text{poly}(n) + O(S(3n))$ size. Running a fast enough CIRCUIT SAT algorithm lets us determine the satisfiability of D in $o(2^n)$ time. Finally, this algorithm for L accepts x iff D is unsatisfiable. To see that this algorithm is correct, observe there is a size- $S(3n)$ circuit W such that D is an unsatisfiable circuit, if and only if there is such a W encoding a satisfying assignment for F_{C_x} , if and only if $x \in L$.

This above argument cannot be carried out directly to prove ACC circuit lower bounds from ACC CIRCUIT SAT algorithms, because of Fact 1. Given an instance x of L , the resulting circuit C_x produced in the reduction from L to SUCCINCT 3SAT can be constructed in polynomial time; however, it looks hard (perhaps impossible) to show that this C_x can be assumed to be an ACC circuit. As C_x is a component of the circuit D , it follows that D itself would not be an ACC circuit, so an ACC CIRCUIT SAT algorithm would not seem to be useful for determining the satisfiability of D .

In the proof of Theorem 3.1 in the author's prior work [Williams 2013], this problem was fixed by settling for a weaker reduction from L to SUCCINCT SAT, which generates an AC^0 circuit C'_x with $3n + O(\log n)$ inputs rather than $n + O(\log n)$. Unfortunately, this constant factor makes a huge difference: to quickly determine satisfiability of the resulting circuit D' in $o(2^n)$ time, a $2^{n/3}/n^{\omega(1)}$ time algorithm for ACC CIRCUIT SAT is needed, instead of a $2^n/n^{\omega(1)}$ algorithm. Algorithms of the former type are not known even for 3SAT; algorithms of the latter type are much more plentiful.

While it is unlikely that these C_x circuits can be implemented in ACC, note that we already assume that ACC is powerful in some sense: in a proof by contradiction, we may assume many functions have small ACC circuits! Since the function computed by C_x is computable in polynomial time, then even if we assume that only P has ACC

circuits, there still exists a circuit C'_x which is ACC and equivalent to C_x , but it is from a nonuniform family, and therefore may be arbitrarily difficult to construct. However, we can use nondeterminism in the algorithm recognizing L in $\text{NTIME}[o(2^n)]$, so at the very least we can guess this elusive C'_x . We also have a good algorithm for ACC CIRCUIT SAT at our disposal. By guessing two more ACC circuits to help us, it turns out that we can always generate a correct ACC circuit C'_x that is equivalent to C_x in $o(2^n)$ time. We arrive at our main lemma.

LEMMA 3.4. *There is a fixed $c > 0$ and $b > 0$ with the following property. Assume P has ACC circuits of depth $d(n)$ and size at most $S(n) \leq 2^{n/4}$. Further assume there is some $k > 2c$ such that ACC CIRCUIT SAT on circuits with $n + k \log n$ inputs, depth $2 \cdot d(3n) + b$, and at most $O(S(3n) + S(2n)n)$ size can be solved in $O(2^n/n^k)$ time.*

Then for every $L \in \text{NTIME}[2^n]$, there is a nondeterministic algorithm A and constant e such that:

- A runs in $O(2^n/n^k + S(3n) \cdot n^e)$ time;
- for every x of length n , $A(x)$ either prints reject or it prints an ACC circuit C'_x with $n + c \log n$ inputs, depth $d(3n)$, and $S(n + c \log n)$ size, such that $x \in L$ if and only if C'_x is the compression of a satisfiable 3-CNF formula of $2^n \cdot n^e$ size; and
- there is always at least one computation path of $A(x)$ that prints a valid circuit C'_x .

That is, given an instance x , the algorithm A nondeterministically generates an equivalent SUCCINCT 3SAT instance C'_x which is an ACC circuit. Let us first give a high-level overview of the proof. Informally, A will guess and verify C'_x in three stages.

- (1) A guesses an ACC circuit D of depth $d(2n)$ and $O(S(2n) \log n)$ size which encodes all the gate and wire information of the circuit C_x which has $cn^c + c$ size. Given a gate index $j = 1, \dots, cn^c + c$, D produces the gate type of j , as well as the indices of gates whose outputs are the inputs for gate j . The correctness of D can be verified in $O(n^c S(2n) \cdot \text{poly}(\log S(2n)))$ time by simply producing the entire circuit of $cn^c + c$ size that is described by D , and checking that it is identical to C_x .
- (2) Next, A guesses an ACC circuit E of depth $d(3n)$ and $O(S(3n))$ size which encodes the outputs of all gates of C_x on all inputs i : given input i and a gate index $j = 1, \dots, cn^c + c$, E produces the output of gate j in C_x evaluated on i . A verifies that E is correct, using the fact that D is correct. By constructing an appropriate ACC CIRCUIT SAT instance that checks for all inputs and all gates that the claimed inputs to that gate are consistent with the output of the gate, this verification takes $O(2^n/n^k)$ time (for k chosen to be greater than $2c$).
- (3) Then, using the fact that E is correct, it is easy to verify that a guessed ACC circuit C'_x is correct via a call to ACC CIRCUIT SAT that runs in $O(2^n/n^k)$ time. A only needs to check if there is an i such that $C'_x(i) \neq E(i, j^*)$, where j^* is the index of the output gate of C_x . (Alternatively, we could just print the circuit $E(\cdot, j^*)$ as a valid ACC circuit that is equivalent to $C_x(\cdot)$.) If E is correct and no such i exists, then C'_x is also correct.

PROOF OF LEMMA 3.4. We describe A in detail. On input x of length n , A guesses an ACC circuit C'_x of size $S(n + c \log n)$, and constructs the SUCCINCT 3SAT circuit C_x with $n + c \log n$ inputs and at most $cn^c + c$ size (of Fact 1) in polynomial time, for a fixed c that is independent of L . By Fact 1, $x \in L$ if and only if C_x is the compression of a satisfiable formula F_{C_x} of $O(2^n n^c)$ length. We must verify that C'_x and C_x compute exactly the same function, using only the algorithm for ACC CIRCUIT SAT.

Without loss of generality, the unrestricted circuit C_x above has gate types AND, OR, NOT, and INPUT, where every AND and OR has fan-in two. By definition, an INPUT

gate has no inputs, and the output value of an INPUT gate is the appropriate input bit itself. The gates are indexed by the numbers $1, \dots, cn^c + c$, where the first $n + c \log n$ indices correspond to the $n + c \log n$ INPUT gates, and the $(cn^c + c)$ th gate is the output gate.

Since the map $x \mapsto C_x$ is polynomial-time computable, the following function f is polynomial-time computable.

Given x , and a gate index $j = 1, \dots, cn^c + c$, $f(x, j)$ outputs the gate type (AND, OR, NOT, INPUT) of the j th gate in the circuit C_x . Furthermore, if the gate type is NOT, then f outputs the gate index j_1 in C_x whose output is the input to j ; if the gate type is an AND or OR, then f outputs the two gate indices j_1 and j_2 in C_x whose outputs are the two inputs of j .

Consider the decision problem D_f :

Given x, j , and $i = 1, \dots, 2c \log n + O(1)$, decide if the i th bit of $f(x, j)$ is 1.

The problem D_f is solvable in polynomial time and hence has $O(S(n + c \log n + O(\log \log n)))$ -size, $d(n + c \log n + O(\log \log n))$ -depth ACC circuits, by assumption.

Let $D(x, j)$ be an ACC circuit implementing the functionality of f . Note we may assume the size of D is

$$O(S(n + O(\log n)) \log n) \leq O(S(2n) \log S(n)), \quad (1)$$

by simply taking $2c \log n + O(1)$ copies of the $S(n + O(\log n))$ -size circuit solving the decision problem D_f . (By convention, let us assume that when D is printing the gate information for an INPUT gate, it prints all-zeroes strings in place of j_1 and j_2 , and when D is printing the information for a NOT gate, it prints all-zeroes in place of j_2 .) Similarly, the depth of D can be assumed to be at most $d(2n)$.

The nondeterministic algorithm \mathcal{A} guesses D , and verifies that D is correct on the given input x in time

$$O(n^c S(n + O(\log n)) \cdot \text{poly}(\log S(n + O(\log n)))) \leq n^c \cdot S(2n) \cdot \text{poly}(\log S(2n)) \leq O(2^{2n/3}),$$

by evaluating $D(x, \cdot)$ on all possible $j = 1, \dots, cn^c + c$, and checking that all outputs of D correspond with the relevant gates in C_x . If D does not output all the gates of C_x correctly, then \mathcal{A} rejects.

Next, consider the following problem.

Given x , an input i of $n + c \log n$ bits, and a gate index $j = 1, \dots, cn^c + c$, output the bit value on the output wire of the j th gate when C_x is evaluated on i .

By assumption, this problem also has ACC circuits, since C_x can be constructed and evaluated on any input i in polynomial time. Let $E(x, i, j)$ be an ACC circuit with this functionality, of size

$$S(n + (n + c \log n) + c \log n + O(1)) \leq S(3n), \quad (2)$$

and depth at most $d(3n)$ as well.

Now, algorithm \mathcal{A} guesses E and wishes to verify its correctness on x . To do this, \mathcal{A} constructs a circuit $\text{VALUE}(i, j)$ built out of D and E , where i has $n + c \log n$ bits and $j = 1, \dots, cn^c + c$. Intuitively, $\text{VALUE}(i, j)$ will output 0 if and only if E produces a sensible output for the j th gate of C_x evaluated on input i .

First, $\text{VALUE}(i, j)$ feeds j to the circuit $D(x, \cdot)$, producing gate indices j_1, j_2 , and a gate type g . VALUE then computes $v_1 = E(x, i, j_1)$, $v_2 = E(x, i, j_2)$ and $v = E(x, i, j)$. (Depending on g , these j_1 and j_2 may be all-zeroes, but this does not matter to us.)

If $g = \text{INPUT}$, then VALUE outputs 0 if and only if $j \in \{1, \dots, n + c \log n\}$ (j is among the first $n + c \log n$ inputs) and the j th bit of i equals v . This behavior can be easily implemented with an AC^0 circuit of size

$$O(n \log n). \quad (3)$$

If $g = \text{NOT}$, then VALUE outputs 0 if and only if $v_1 = \neg v$.

If $g = \text{AND}$, then VALUE outputs 0 if and only if $v_1 \wedge v_2 = v$.

If $g = \text{OR}$, then VALUE outputs 0 if and only if $v_1 \vee v_2 = v$.

Note that each of these three conditions can be implemented with a constant number of gates, given the values g , v_1 , v_2 , and v . It follows that VALUE can be implemented as an ACC circuit.

Since \mathcal{A} has not rejected, D is correct, so we know that for all i, j , the gate types g and input connections j_1 and j_2 are correct. Therefore, $\text{VALUE}(i, j) = 1$ if and only if E asserts that the output of gate j in $C_x(i)$ equals v , and E asserts the inputs to j have values v_1, v_2 , but the gate type g dictates that the output of j should be $\neg v$. It follows that VALUE is an unsatisfiable circuit if and only if E prints correct values for all gates in $C_x(i)$, over all i .

Therefore, by calling ACC circuit satisfiability on $\text{VALUE}(\cdot, \cdot)$, \mathcal{A} determines whether E is correct. The algorithm \mathcal{A} rejects if E is deemed incorrect. The circuit $\text{VALUE}(i, j)$ has $n + 2c \log n + O(1)$ inputs, depth at most $d(2n) + d(3n) + b$ for a fixed constant b , and it follows from (1), (2), and (3) that the circuit for VALUE has $O(S(3n) + S(2n) \log S(2n) + n \log S(n)) \leq O(S(3n) + S(2n)n)$ size. By assumption, the assumed ACC satisfiability algorithm runs in $O(2^n/n^k)$ time for some $k > 2c$.

After checking that E is a correct guess, the question of whether C'_x is equivalent to C_x can now be verified. (Alternatively, at this point we may simply print the circuit $E(\cdot, cn^c + c)$ as a valid circuit that is equivalent to $C_x(\cdot)$.) First, note that if E is correct, then for all i , $C_x(i) = E(x, i, cn^d + c)$. Therefore, it suffices to set up an ACC circuit $\text{EQUIV}(i)$ which outputs 1 if and only if $C'_x(i) \neq E(x, i, cn^c + c)$, and determine if EQUIV is satisfiable using the algorithm for ACC CIRCUIT SAT. Since $\text{EQUIV}(i)$ has $n + c \log n$ inputs, depth $d(3n) + O(1)$, and size $O(S(n + O(\log n)))$, the circuit satisfiability call runs in $O(2^n/n^k)$ time, by assumption. If EQUIV is satisfiable, then \mathcal{A} rejects.

Finally, \mathcal{A} prints its guessed circuit C'_x if the algorithm did not reject on any of the above steps. \square

Remark 3.5. The proof of the lemma does not require specific properties of ACC. We only need that the underlying circuit class \mathcal{C} contains AC^0 and is closed under composition of two circuit families. The same goes for the proof of Theorem 3.2.

Remark 3.6. In fact, the lemma shows that, given any circuit C and a \mathcal{C} -circuit D , we can efficiently check if C is equivalent to D using nondeterminism (under the assumptions that \mathcal{P} has \mathcal{C} -circuits and there are efficient \mathcal{C} -SAT algorithms). If \mathcal{P} has \mathcal{C} -circuits, then Circuit Evaluation has \mathcal{C} -circuits. Therefore, any circuit family $\{C_n\}$ will have an equivalent \mathcal{C} -circuit family, as well as \mathcal{C} -circuits encoding the “gate information” and “gate values” of $\{C\}$.

With Lemma 3.4 in hand, the proof of Theorem 3.2 closely follows the author’s prior work (Theorem 3.1), except the circuit C'_x is substituted in place of C_x . Let us give the details. We use the specific example of ACC in place of a generic circuit class \mathcal{C} ; modification for other classes \mathcal{C} is straightforward.

REMINDER OF THEOREM 3.2. *Let $S(n) \leq 2^{n/4}$. There is a $c > 0$ and $b > 0$ such that, if \mathcal{C} -CIRCUIT SAT instances with at most $n + c \log n$ variables, depth at most*

$2 \cdot d(3n) + b$, and $O(n \cdot S(2n) + S(3n))$ size can be solved in $O(2^n/n^c)$ time, then E^{NP} does not have nonuniform C circuits of depth $d(n)$ and $S(n)$ size.

PROOF OF THEOREM 3.2. Suppose ACC CIRCUIT SAT instances with $n + c \log n$ variables, depth $2 \cdot d(3n) + O(1)$, and $O(n \cdot S(2n) + S(3n))$ size can be solved in $O(2^n/n^c)$ time for a sufficiently large c . Further, suppose that E^{NP} has nonuniform ACC circuits of depth d and $S(n)$ size. The goal is to show that $NTIME[2^n] \subseteq NTIME[o(2^n)]$, contradicting the nondeterministic time hierarchy [Seiferas et al. 1978; Žák 1983].

Let $L \in NTIME[2^n]$. We describe a fast nondeterministic algorithm \mathcal{B} deciding L . As discussed earlier (Theorem 2.1), we may assume L has a multitape Turing machine implementation in $O(2^n)$ time, and we only need to simulate L on a RAM in $O(2^n/n^c)$ time for large enough c to obtain the contradiction.

On input x of length n , \mathcal{B} first runs the nondeterministic algorithm \mathcal{A} of Lemma 3.4. Using the ACC CIRCUIT SAT algorithm and the fact that P has ACC circuits, \mathcal{A} runs in $O(2^n/n^c + S(3n) \cdot \text{poly}(n)) \leq O(2^n/n^c)$ time, and for some computation path, \mathcal{A} produces an ACC circuit C'_x of $S(n + c \log n)$ size, $d(n + c \log n)$ depth, and $n + c \log n$ inputs such that $x \in L$ if and only if C'_x is the compression of a satisfiable formula $F_{C'_x}$.

Then, \mathcal{B} nondeterministically guesses a $S(3n)$ -size $d(3n)$ -depth circuit W . By Fact 2, there exists such a W that encodes a satisfying assignment for $F_{C'_x}$ if and only if $x \in L$.

Next, \mathcal{B} constructs an ACC CIRCUIT SAT instance D to verify that W is correct (just as in the proof of Theorem 3.1). The circuit D has $n + c \log n$ inputs fed to $O(n)$ copies of C'_x , so that when i is input to D , the i th clause of the 3CNF formula $F_{C'_x}$ is printed on $O(n)$ bits of output. The $O(n)$ bits encode three variable indices along with sign bits for each variable. For the three variables, an assignment is computed for them by evaluating the indices on three copies of W . Finally, D compares the sign bits with the bits output by the copies of W , and outputs 0 iff the variable assignment encoded by W satisfies the i th clause. Now observe that our guessed W encodes a satisfying assignment for $F_{C'_x}$, if and only if every clause of $F_{C'_x}$ is satisfied by the assignment encoded by W , if and only if $D(i) = 0$ for all i , if and only if D is unsatisfiable.

Observe that D has $O(n \cdot S(2n) + S(3n))$ size, depth at most $2d(3n) + O(1)$, and $n + c \log n$ inputs. By assumption, the satisfiability of D can be determined in $O(2^n/n^c)$ time, hence \mathcal{B} decides if $x \in L$ in $O(2^n/n^c)$ time. \square

4. A SATISFIABILITY ALGORITHM FOR ACC CIRCUITS

Now, we present an algorithm that determines the satisfiability of ACC circuits slightly faster than the 2^n runtime of exhaustive search. There are two components in the algorithm: a nice representation of ACC circuits, and a method for evaluating this representation quickly on all of its inputs. This method can be implemented using either fast rectangular matrix multiplication, a dynamic programming approach, or a divide-and-conquer recursive approach.

It follows from the work of Yao [1990], Beigel and Tarui [1994], and Allender and Gore [1991] that, given any ACC circuit of size s , one can produce a $s^{O(\log^c s)}$ size SYM^+ circuit in $\text{poly}(s^{O(\log^c s)})$ time that has equivalent functionality. (For more background, see the Preliminaries.)

LEMMA 4.1. *There is an algorithm and function $f : \mathbb{N} \times \mathbb{N} \rightarrow \mathbb{N}$ such that given an ACC circuit with MOD_m gates of depth d and size s , the algorithm outputs an equivalent SYM^+ circuit of $s^{O(\log^{f(d,m)} s)}$ size, with ANDs of $O(\log^{f(d,m)} s)$ fan-in. The algorithm takes at most $s^{O(\log^{f(d,m)} s)}$ time.*

Furthermore, given the number of ANDs in the circuit that evaluate to 1, the symmetric function itself can be evaluated in $s^{O(\log^{f(d,m)} s)}$ time.

The function $f(d, m)$ is estimated to be no more than $m^{O(d)}$. Technically speaking, Lemma 4.1 is not explicitly proved in prior work, but Allender and Gore effectively show it: they prove that given a *uniform* ACC circuit (with an efficiently computable connection language), there is a similarly *uniform* SYM^+ circuit of the appropriate size. Their proof corresponds to an efficient, deterministic algorithm computing the transformation, and this algorithm works equally well if it is simply given any ACC circuit as input (not necessarily uniform). Since the proof of the lemma is essentially identical to previous proofs already in the literature, it has been placed in Appendix A.

We stress that the transformation from ACC to SYM^+ results in a circuit with *no negated variables*: just a symmetric function of ANDs of variables. (This property is used in the dynamic programming algorithm here.) It is a simple exercise to transform a SYM of ANDs of polylog fan-in (with negations on some of the wires coming out of input variables) to a SYM of ANDs of variables, while preserving the relevant parameters of Lemma 4.1. A proof sketch of this transformation is in Appendix B.

4.1. Rapid Evaluation of an ACC Circuit on All of Its Inputs

The other component of the ACC satisfiability algorithm is a method for rapidly evaluating a given SYM^+ circuit on all of its possible satisfying assignments.

LEMMA 4.2 (EVALUATION LEMMA). *There is an algorithm that, given a SYM^+ circuit of size $s \leq 2^{1n}$ and n inputs with a symmetric function that can be evaluated in $\text{poly}(s)$ time, runs in $(2^n + \text{poly}(s)) \cdot \text{poly}(n)$ time and prints a 2^n -bit vector V which is the truth table of the function represented by the given circuit. That is, $V[i] = 1$ iff the SYM^+ circuit outputs 1 on the i th variable assignment.*

That is, any SYM^+ circuit can be evaluated on all 2^n assignments in *polynomial amortized time per assignment*. Brute force search would take $2^n \cdot \text{poly}(s)$ time, but the algorithm manages to use roughly $2^n + \text{poly}(s)$ time instead.

Lemma 4.2 can be proved in three different ways; each are appealing for different reasons. The first proof, using dynamic programming, has the benefit that it can be completely described with few technical details. For pedagogical purposes, it is preferred. The second proof uses a powerful primitive (fast matrix multiplication) that is common in theoretical computer science, and it is plausible that the matrix multiplication approach could be extended further. The third proof uses a simple divide-and-conquer approach: we can reduce the SYM^+ circuit evaluation problem to that of evaluating a multilinear n -variable polynomial on all 2^n assignments over $\{0, 1\}^n$, and solve that problem in $O(2^n \cdot \text{poly}(n))$ time. (This algorithm is discussed in the author's survey article [Williams 2011], and is omitted here.)

Proof 1. Dynamic Programming. The evaluation lemma can be proved using simple dynamic programming, following a conversation with Andreas Björklund.

PROOF OF LEMMA 4.2. Assume we are given a SYM^+ circuit C'' with a collection of s'' AND gates over some variables $\{x_1, \dots, x_n\}$. Let $G_j \subseteq [n]$ be the set of variable indices that are input to the j th AND gate. Define a function $f : 2^{[n]} \rightarrow \mathbb{N}$, where $f(S)$ equals the number of $j = 1, \dots, s''$ such that $S = G_j$. The function f can be prepared as a lookup table in $O(2^n + s'' \cdot \text{poly}(n)) \leq O(2^n)$ time, by building a table of 2^n entries which are initially zero, and for each of the s'' AND gates corresponding to a subset S , we increment the S th entry in the table.

Now consider the function $g(T) = \sum_{S \subseteq T} f(S)$ defined on all $T \subseteq [n]$. (Typically, g is called the *zeta transform of f* .) Observe that $g(T)$ equals the number of AND gates set to 1 on the variable assignment obtained by setting $x_i = 1$ for $i \in T$, and $x_i = 0$ for $i \notin T$. Therefore, the table of 2^n integers of size $O(\log s'')$ representing the function g is equivalent to the matrix N in the previous proof. Hence, if we can compute g , then we can evaluate C'' on all of its possible inputs.

It remains to show how to compute g efficiently. Given f , the function g can be computed in $O(2^n \cdot \text{poly}(n))$ time by a dynamic programming algorithm of Yates from 1937 (cf. Björklund et al. [2009, Section 2.2]). For $i = 0, \dots, n$, define $g_i : 2^{[n]} \rightarrow \mathbb{N}$ by $g_0(T) = f(T)$, and

$$g_i(T) = \begin{cases} g_{i-1}(T) + g_{i-1}(T \setminus \{i\}) & \text{if } i \in T, \\ g_{i-1}(T) & \text{otherwise.} \end{cases}$$

It follows that each g_{i+1} can be obtained from g_i in $O(2^n \cdot \text{poly}(n))$ time. Induction shows that $g_i(T) = \sum_{S \subseteq T} f(S)$ where the sum is over all $S \subseteq T$ subject to the condition that $\{j \in S \mid j > i\} = \{j \in T \mid j > i\}$. When $i = n$, both of these sets are always empty, so it follows that $g_n = g$. \square

This description is suitable for random access machines, but the algorithm can also be implemented on a multitape Turing machine using standard ideas. (Strictly speaking, a multitape implementation is not necessary to prove ACC lower bounds, because Lemma 2.1 shows it suffices to have a fast random-access implementation of any $L \in \text{NTIME}_{\text{TM}}[2^n]$. However, the extension to multitape may be useful for future work.)

Proof 2. Matrix Multiplication. An alternative way to prove Lemma 4.2 is to use a fast rectangular matrix multiplication algorithm of Coppersmith, building on prior work of Schönhage [1981]. This algorithm works in the case where the “middle” dimension of the matrices is polynomially smaller than the other two. In this case, matrix multiplication can be done nearly optimally.⁵

LEMMA 4.3 [COPPERSMITH 1982]. *For all sufficiently large N , multiplication of an $N \times N^1$ matrix with an $N^1 \times N$ matrix can be done in $O(N^2 \log^2 N)$ arithmetic operations.*

More precisely, Coppersmith shows that there is a constant K such that one can multiply an $N \times N$ matrix with an $N \times N^1$ matrix in $K \cdot N^2 \log^2 N$ operations with a bilinear algorithm, that is, a depth-3 arithmetic circuit with additions at the top level, multiplications in the middle, and additions at the bottom level, where each input wire to an addition gate may also multiply the input by a scalar. From the duality of bilinear matrix multiplication algorithms [Hopcroft and Musinski 1973], a bilinear algorithm for multiplying $N \times N$ and $N \times M$ directly implies a bilinear algorithm for multiplying $N \times M$ and $M \times N$. Furthermore, Coppersmith’s algorithm is explicit, in that it can be executed on typical machine model (even a multitape TM) in $O(N^2 \cdot \text{poly}(\log N))$ time, on matrices over any field of $\text{poly}(\log N)$ elements. A discussion of implementation details for his construction can be found in Appendix C. For us, the relevant corollary is the following.

⁵Curiously, later work on rectangular matrix multiplication from the 90’s [Coppersmith 1997; Huang and Pan 1998] does not provide tight enough bounds: only $N^{2+\varepsilon}$ for all $\varepsilon > 0$, rather than $N^2 \log^2 N$. Note that a bound of $N^2 \cdot 2^{(\log N)^{o(1)}}$ already suffices for our application.

COROLLARY 4.4. *For all sufficiently large N , two 0 - 1 matrices of dimensions $N \times N^{-1}$ and $N^{-1} \times N$ can be multiplied over the integers in $O(N^2 \cdot \text{poly}(\log N))$ time.*

We arrive at our second proof of Lemma 4.2.

PROOF OF LEMMA 4.2. Suppose we are given a SYM^+ circuit C'' of size $s'' \leq 2^{1n}$. Partition the inputs of C'' into two sets A and B of size at most $n' = (n + 1)/2$ each. Set up two matrices M_A and M_B of dimensions $2^{n'} \times s''$ and $s'' \times 2^{n'}$ (respectively). The rows of M_A are indexed by all possible assignments to the variables in set A , while the columns of M_A are indexed by the AND gates of C'' . Similarly, the columns of M_B are indexed by variable assignments in B , while the rows of M_B are indexed by the ANDs of C'' . Define:

$$M_A(i, j) = \begin{cases} 1 & \text{if the } i\text{th variable assignment in } A \text{ doesn't force the } j\text{th AND to be } 0, \\ 0 & \text{otherwise,} \end{cases}$$

and

$$M_B(j, k) = \begin{cases} 1 & \text{if the } k\text{th variable assignment in } B \text{ doesn't force the } j\text{th AND to be } 0, \\ 0 & \text{otherwise.} \end{cases}$$

Note the preparation of M_A and M_B takes at most $2^{n/2} \cdot s'' \cdot \text{poly}(n) \leq O(2^{n/2+2n})$ time.

Multiply M_A and M_B , yielding a matrix N . Note that $M_A(i, j) \cdot M_B(j, k) = 1$ iff the i th assignment in A and the k th assignment in B together set the j th AND of C'' to 1. (Given an assignment to all variables in A and B , the AND is forced to either 1 or 0.) Hence, $N(i, k)$ equals the number of ANDs set to 1 by the i th assignment in A and the k th assignment in B . Therefore, C'' is satisfiable if and only if some entry of N makes the symmetric function of C'' output 1.

Since $s'' \leq 2^{1n'}$, the fast rectangular matrix multiplication of Corollary 4.4 applies, and the multiplication of M_A and M_B can be done in $2^{2n'} \text{poly}(n)$ time.

To decide whether some entry of N makes the symmetric function output 1, initialize a bit vector T of length $s'' + 1$, setting $T[i]$ to be the value of the symmetric function of C'' on each $i = 0, 1, \dots, s''$. The construction of T takes $\text{poly}(s'')$ time, since the symmetric function can be evaluated in $\text{poly}(s'')$ time. Then, for every pair $i, k \in \{1, \dots, 2^{n'}\}$, if $v[N(i, k)] = 1$, then stop and report *satisfiable*. If every pair has been examined without stopping, report *unsatisfiable*. The for-loop over all pairs can be implemented in time $2^{2n'} \cdot \text{poly}(n) \leq 2^n \cdot \text{poly}(n)$ by standard table lookup or by sorting the distinct elements of $N(i, k)$. \square

4.2. The Final Algorithm

Given the evaluation lemma, the ACC satisfiability algorithm is relatively straightforward.

THEOREM 4.5. *For every $d > 1$ and $m > 1$, there is an $\varepsilon \in (0, 1)$ such that satisfiability of depth- d ACC circuits with MOD_m gates, n inputs, and 2^{n^ε} size can be determined in $2^{n - \Omega(n^\delta)}$ time for some $\delta > \varepsilon$ that depends only on d and m .*

PROOF. Let ℓ, ε be parameters to set later. Suppose we are given a depth- d ACC circuit C of $s = 2^{n^\varepsilon}$ size and n inputs. Make a circuit C' with $s \cdot 2^\ell$ size and $n - \ell$ inputs which is obtained by producing 2^ℓ copies of C , plugging in a different possible assignment to the first ℓ inputs of C in each copy, and taking the OR of these copies. Observe C' is a depth- $(d + 1)$ ACC circuit, and C is satisfiable if and only if C' is satisfiable.

Applying the translation from ACC to SYM^+ (Lemma 4.1), a circuit C'' can be produced which is equivalent to C' , where C'' consists of a symmetric gate connected to $s'' \leq s^{e(\ell^e \log^e s)}$ ANDs of variables, for some constant e that depends on the depth d and modulus m . Producing C'' from C' takes only $s^{O(\ell^e \log^e s)}$ steps. When $s = 2^{n^\varepsilon}$, $s'' \leq 2^{en^\varepsilon(\ell^e n^{\varepsilon e})}$. Set $\ell = n^{1/(2e)}$, and observe that $s'' \leq 2^{n^{2/3}}$ for all sufficiently large n and sufficiently small ε .

By the evaluation lemma (Lemma 4.2) and the fact that the symmetric function of C'' can be evaluated in $\text{poly}(s'')$ time, C'' can be evaluated on all of its possible assignments in $O(2^{n-\ell} \cdot \text{poly}(n)) \leq 2^{n-\Omega(n^{1/(2e)})}$ time; hence, the satisfiability of C can be determined within this time. \square

Two Remarks. It is worth pointing out a couple more things about the algorithm. First, the algorithm can be generalized in multiple ways which may be useful in the future. Instead of taking an OR of all partial assignments to a small number of variables in C , one could instead take any constant number of ANDs and ORs of partial assignments, convert this to a SYM^+ circuit, then apply the evaluation lemma. This observation shows that any quantified Boolean formula with a constant number of quantifier blocks and a predicate described by an ACC circuit of subexponential size can also be solved faster than exhaustive search. Second, note that the algorithm does not give a faster way to solve satisfiability for the class SYM^+ itself, because in the algorithm we need that the OR of 2^ℓ circuits from the class is still a circuit in the class. Hence, we cannot give lower bounds for SYM^+ at the present time.

5. ACC LOWER BOUNDS

Combining the results of the previous two sections, nonuniform lower bounds for ACC can be proved.

REMINDER OF THEOREM 1.2. *For every d and m , there is a $\delta > 0$ and a language in E^{NP} that fails to have nonuniform ACC circuits of depth d and size 2^{n^δ} with MOD_m gates.*

PROOF. Theorem 4.5 states that for every d and m there is an $\varepsilon > 0$ so that satisfiability of depth- d ACC circuits with MOD_m gates, n inputs, and $2^{O(n^\varepsilon)}$ size can be solved in $2^{n-\Omega(n^\delta)}$ time, for some $\delta > \varepsilon$. Theorem 3.2 says there is a $c > 0$ such that, if ACC CIRCUIT SAT instances with $n + c \log n$ variables, depth $2d + O(1)$, and at most $s = n \cdot 2^{O(n^\varepsilon)}$ size can be solved in $O(2^n/n^c)$ time, then E^{NP} does not have non-uniform ACC circuits of depth d and 2^{n^ε} size. The lower bound follows, as $2^{(n+c \log n)-\Omega((n+c \log n)^\delta)} \ll O(2^n/n^c)$ for every c . \square

It follows that complete problems such as SMALLEST SUCCINCT 3SAT (given a circuit C and integer i , output the i th bit of the smallest satisfying assignment to the formula F_C encoded by C) require exponential ACC circuits. The E^{NP} lower bound can be “padded down” in a standard way to prove superpolynomial lower bounds for a class that is very close to P^{NP} .

COROLLARY 5.1. *For every d and m , $\text{QuasiP}^{\text{NP}} = \text{TIME}[n^{\log^{O(1)} n}]^{\text{NP}}$ does not have nonuniform ACC circuits of depth d , MOD_m gates, and polynomial size.*

PROOF. If there were d and m such that $\text{TIME}[2^{(\log n)^c}]^{\text{NP}}$ had such circuits for every c , then by a padding argument (replacing n with $2^{n^{1/c}}$), it would follow that E^{NP} has depth- d size- $2^{O(n^{1/c})}$ circuits for every c , contradicting Theorem 1.2. \square

Note it is known that $\text{NTIME}[n^{\log^{O(1)} n}]^{\text{NP}}$ does not have polynomial size (unrestricted) circuits [Kannan 1982].

Superpolynomial ACC lower bounds for NEXP are also provable. First, we need a theorem established in prior work: if NEXP has (unrestricted) polynomial-size circuits, then every satisfiable formula output by the SUCCINCT 3SAT reduction in Fact 1 has some *satisfying assignment* that can be represented with a polynomial-size unrestricted circuit.

More precisely, say that SUCCINCT 3SAT *has succinct satisfying assignments* if there is a fixed constant c such that for every language $L \in \text{NTIME}[2^n]$ and every $x \in L$ of length n , there is a circuit W_x of $\text{poly}(n)$ size with $k \leq n + c \log n$ inputs such that the variable assignment $z_i = W(i)$ for all $i = 1, \dots, 2^k$ is a satisfying assignment for the formula F_{C_x} , where C_x is the circuit obtained by the SUCCINCT 3SAT reduction in Fact 1. Say that W_x is a *succinct satisfying assignment* for C_x .

THEOREM 5.2 [WILLIAMS 2013]. *Suppose NEXP has polynomial-size circuits. Then, SUCCINCT 3SAT has succinct satisfying assignments.*

Theorem 5.2 is not explicitly proved in the article, however, it follows immediately from another theorem. Say that NEXP has *universal witness circuits of polynomial size* if for every $L \in \text{NEXP}$ and every correct exponential time verifier for L , there is a $c > 0$ such that for every $x \in L$, there is a circuit of size at most $|x|^c + c$ which encodes a witness for x that is accepted by the verifier. (For more formal definitions, see Williams [2013].) The following directly implies Theorem 5.2.

THEOREM 5.3 [IMPAGLIAZZO ET AL. 2002; WILLIAMS 2013]. *If $\text{NEXP} \subseteq \text{P/poly}$, then every language in NEXP has universal witness circuits of polynomial size.*

The proof of Theorem 5.3 follows an argument by Impagliazzo et al. [2002].

The second ingredient in the lower bound for NEXP is a simple folklore lemma.

LEMMA 5.4 FOLKLORE. *Let \mathcal{C} be any circuit class. If P has nonuniform \mathcal{C} circuits of $S(n)^{O(1)}$ size, then there is a $c > 0$ such that every $T(n)$ -size circuit family (uniform or not) has an equivalent $S(n + O(T(n) \log T(n)))^c$ -size circuit family in \mathcal{C} .*

PROOF. If P has nonuniform $S(n)^{O(1)}$ -size \mathcal{C} circuits, then for some $c > 0$, the CIRCUIT EVAL problem has $S(n)^c$ -size circuits. (Recall the CIRCUIT EVAL problem is: given an arbitrary Boolean circuit C and input x , evaluate C on x and output the answer.) Let $\{D_n(\cdot, \cdot)\}$ be a $S(n)^c$ -size circuit family for this problem. Now let $\{C_n\}$ be an arbitrary $T(n)$ -size circuit family. To obtain an equivalent \mathcal{C} -circuit family $\{C'_n\}$ of $S(n + O(T(n) \log T(n)))^c$ size, define $C'_{|x|}(x) = D_{n_1}(C_{|x|}, x)$ for an appropriate length $n_1 \leq n + O(T(n) \log T(n))$. \square

Note if $S(n)$ and $T(n)$ are polynomials, then $S(n + O(T(n) \log T(n)))^c$ is also polynomial. Also note Lemma 5.4 implies that, if $\text{P} \subseteq \text{ACC}$, then there are *fixed constants* d and m such that every polynomial-size circuit family has equivalent polynomial-size depth- d circuits with AND, OR, NOT, and MOD_m gates. (Here, d and m are the constants arising in the resulting ACC circuit family for CIRCUIT EVAL.)

REMINDER OF THEOREM 1.1. $\text{NTIME}[2^n]$ *does not have nonuniform ACC circuits of polynomial size.*

PROOF. First, we claim that if $\text{NTIME}[2^n]$ has polysize ACC circuits, then every language in NEXP has polysize ACC circuits. Let us sketch this implication, for completeness. If $\text{NTIME}[2^n]$ has polysize ACC circuits, then the NEXP-complete problem SUCCINCT BOUNDED HALTING has polysize ACC circuits: Given a nondeterministic

machine N , string x , and t written in binary, does $N(x)$ have an accepting computation path of length at most t ? The reduction from any $L \in \text{NEXP}$ to $\text{SUCCINCT BOUNDED HALTING}$ can be expressed with an AC^0 circuit of size $\text{poly}(n, \log t)$. (Take any nondeterministic machine N with running time 2^{n^k} that accepts L . Given an input x , the AC^0 circuit outputs the code of N as the first input of the $\text{SUCCINCT BOUNDED HALTING}$ instance, x as the second input, and $2^{|x|^k}$ as the third input, written in binary. This only needs an AC^0 circuit that outputs 1 followed by $|x|^k - 1$ zeroes.) Hence, every $L \in \text{NEXP}$ can be recognized by an ACC circuit family of size n^ℓ , for some ℓ depending on L .

By Lemma 5.4 and Theorem 5.2, it follows that SUCCINCT 3SAT has succinct satisfying assignments that are polynomial-size ACC circuits. We claim that a contradiction can be obtained by carefully examining the proof of Theorem 1.2 (the lower bound for E^{NP}). There, the only place requiring the full assumption “ E^{NP} has nonuniform ACC circuits of size $S(n)$ ” is inside the proof of Theorem 3.2. In particular, the assumption is needed in Fact 2, where it is shown that for every satisfiable instance of SUCCINCT 3SAT , at least one of its satisfying assignments can be encoded in a size- $S(3n)$ ACC circuit. (The only other part of Theorem 3.2 where the assumption is applied is Lemma 3.4, but there it is only required that P has nonuniform ACC circuits.) But from this proof, we already have that SUCCINCT 3SAT has succinct satisfying assignments which are ACC circuits.

Hence, the ACC CIRCUIT SAT instance D constructed in Theorem 3.2 with the witness circuit W has size polynomial in its $n + c \log n$ inputs. Finally, the Circuit SAT algorithm of Theorem 4.5 can determine satisfiability of any $n + c \log n$ input, n^c size ACC circuit in $O(2^{n - \log^2 n})$ time, for every constant c . Therefore unsatisfiability of D can be determined in $O(2^n / n^c)$ time for every constant c , and the desired contradiction follows from the nondeterministic time hierarchy. \square

It follows that problems complete under AC^0 reductions for NEXP such as SUCCINCT 3SAT (given a circuit C , does it encode a satisfiable 3-CNF formula F_C ?) require superpolynomial-size ACC circuits.

5.1. An Extension to “Half-Exponential” Type Lower Bounds

The NEXP lower bounds can be extended a little by studying the proof of Theorem 5.3. However, the results are a bit ugly, so let us only sketch the arguments. A function $f : \mathbb{N} \rightarrow \mathbb{N}$ is said to be *sub-half-exponential* if for every k , $f(f(n^k)^k) \leq 2^{n^{o(1)}}$. The following was conjectured by Russell Impagliazzo (private communication), and can be proved by augmenting Theorem 5.3 with other known results.

THEOREM 5.5. *Let $S(n)$ be any sub-half-exponential function such that $S(n) \geq n$ for all n . If $\text{NTIME}[2^n]$ has $S(n)$ size circuits, then all languages in NEXP have universal witness circuits of size $O(S(S(n^c)^c)^c)$, for some c depending on the language.*

The proof goes along the lines of Theorem 5.3, but with $S(n)$ substituted in place of polynomials: we assume (a) NEXP does not have universal witness circuits of $S(S(n^c)^c)^c$ size for any c , (b) $\text{NTIME}[2^n]$ does have $S(n)$ circuits, and derive a contradiction from the two. Assumption (a) implies that in time $O(2^n)$, one can nondeterministically guess and verify the truth table of a Boolean function on n bits that requires $S(S(n^c)^c)^c$ size circuits for every c , for infinitely many inputs. This is enough to partially derandomize $\text{MATIME}[S(S(n^{O(1)})^{O(1)})^{O(1)}]$ for infinitely many input lengths, putting the class inside of i.o.- $\text{NTIME}[2^n]/n$ [Impagliazzo et al. 2002] (recall that $S(S(n^{O(1)})^{O(1)})^{O(1)} \leq 2^{n^{o(1)}}$). Assumption (b) implies that NEXP has $S(n^{O(1)})^{O(1)}$ circuits, hence $\text{TIME}[2^{S(n^{O(1)})^{O(1)}}] \subseteq \text{MATIME}[S(S(n^{O(1)})^{O(1)})^{O(1)}]$ [Babai et al. 1993; Miltersen

et al. 1999]. It also follows from assumption (b) that i.o.-NTIME[2^n] / n has $O(S(n))$ size circuits on infinitely many input lengths. Putting these containments together, it follows that TIME[$2^{S(n)^{O(1)}}$] has $O(S(n))$ size circuits on infinitely many input lengths. This is false by direct diagonalization: for all large enough n there is a function f on n variables with circuit complexity greater than $S(n)^2$, and the lexicographically first f can be found in $2^{O(S(n)^3)}$ time and simulated on a given input.

Combining Theorem 5.5 and Lemma 5.4, we immediately obtain the following implication between ACC circuits for NTIME[2^n] and ACC circuits which encode witnesses for NEXP.

COROLLARY 5.6. *If NTIME[2^n] has $S(n)$ -size ACC circuits, then every language in NEXP has universal witness ACC circuits of $S(S(S(n^c)^c)^c)^c$ for some c depending on the language.*

One extra application of S comes from Theorem 5.5, which produces universal witness circuits; the other comes from Lemma 5.4, which converts those circuits to ACC. Define $f : \mathbb{N} \rightarrow \mathbb{N}$ to be *sub-third-exponential* if for every k , $f(f(n^k)^k)^k \leq 2^{n^{O(1)}}$. Examples of sub-third-exponential functions are $f(n) = n^{\text{poly}(\log n)}$ and $f(n) = 2^{2^{\text{poly}(\log \log n)}}$.

THEOREM 5.7. *NTIME[2^n] does not have sub-third-exponential size ACC circuits.*

The argument is the same as Theorem 1.1, except we apply Corollary 5.6: if NTIME[2^n] has such circuits, then Corollary 5.6 says that NEXP has universal witness circuits which are ACC and have subexponential size. This implies that SUCCINCT 3SAT instances have subexponential size ACC circuits that encode their satisfying assignments, which is enough to establish the contradiction in Theorem 1.1.

THEOREM 5.8. *Let $g : \mathbb{N} \rightarrow \mathbb{N}$ have the property that there is a sub-third-exponential function f satisfying $g(f(n)) \geq 2^n$. Then, NTIME[$g(n)$] does not have polynomial-size ACC circuits.*

If such circuits did exist, then by padding, NTIME[2^n] \subseteq NTIME[$g(f(n))$] would have ACC circuits of size $f(n)^{O(1)}$ for some sub-third-exponential f , contradicting Theorem 5.7. (Raising f to a constant power is still a sub-third-exponential function.) It follows that the polynomial-size lower bound can be extended down to grotesque classes such as NTIME[$2^{2^{\sqrt{\log \log n}}}] \subsetneq$ NTIME[2^n], since $f(n) = 2^{(\log n)^{\log \log n}} = 2^{2^{(\log \log n)^2}}$ is sub-third-exponential, and $g(f(n)) \geq 2^n$ for functions like $g(n) = 2^{2^{\sqrt{\log \log n}}}$.

Finally, it is also straightforward to extend the lower bounds to polysize ACC circuits of slightly nonconstant depth, as the ACC SAT algorithm still beats exhaustive search on polynomial-size circuits of depth $o(\log \log n)$ with constant moduli. The details can be found by studying Theorem 4.5 or by reading Wang [2011].

6. CONCLUSION

This article demonstrates that the research program of proving circuit lower bounds via satisfiability algorithms is a viable one. Further work will surely improve the results. Three natural next steps are: replace ACC with TC⁰ circuits in the lower bounds, or replace NEXP with EXP, or extend the exponential lower bounds from E^{NP} to NEXP.

The results of Section 3 and Lemma 5.4 show that one only has to find a very minor improvement in algorithms for TC⁰ satisfiability in order to establish nonuniform TC⁰ lower bounds for NEXP. The author sees no serious impediment to the existence of such an algorithm; he can only report that the algorithms tried so far do not work. The evaluation lemma for SYM⁺ circuits is key to the ACC SAT algorithm, and it would be very

interesting to find similar lemmas for TC^0 or NC^1 . It is plausible that the characterization of NC^1 as bounded-width branching programs [Barrington 1989] could be applied to prove an analogous evaluation lemma for Boolean formulas, which would lead to nontrivial depth lower bounds for NEXP. (Note that permutation branching programs of width 4 can be simulated in ACC [Barrington and Thérien 1988], while width 5 captures NC^1 .) Along the lines of the author's prior work [Williams 2013], Oded Goldreich and Or Meir (personal communication) have observed that the consequence of Theorem 3.2 holds even when we replace \mathcal{C} -CIRCUIT SAT with the problem: given an n -input $S(n)$ -size \mathcal{C} -circuit, approximate its probability of acceptance on a uniform random input to within a $1/6$ additive factor. It is widely believed that this problem can be solved in *polynomial time* for any reasonable \mathcal{C} , so the hypothesis appears easier to satisfy. Recent work with Santhanam has further weakened the derandomization hypothesis necessary for circuit lower bounds [Santhanam and Williams 2013].

It should be possible to extend the superpolynomial lower bound for ACC down to the class $QuasiNP = NTIME[n^{\log^{O(1)} n}]$. This article comes fairly close to proving this result. The only step missing is a proof of the implication: "if QuasiNP has polynomial-size ACC circuits, then there are polynomial-size ACC circuits that encode witnesses to QuasiNP languages." A couple of lemmas rely only on P having nonuniform ACC circuits, so they could be potentially applied in proofs of even stronger lower bounds. At any rate, the prospects for future circuit lower bounds look very promising.

APPENDIXES

A. PROOF OF LEMMA 4.1

REMINDER OF LEMMA 4.1. *Let m and d be fixed constants. There is an algorithm and function $f : \mathbb{N} \times \mathbb{N} \rightarrow \mathbb{N}$ such that given an ACC circuit with MOD_m gates of depth d and size s , the algorithm outputs an equivalent SYM^+ circuit of $s^{O(\log^{f(d,m)} s)}$ size. The algorithm takes at most $s^{O(\log^{f(d,m)} s)}$ time.*

Furthermore, given the number of ANDs in the circuit that evaluate to 1, the symmetric function itself can be evaluated in $s^{O(\log^{f(d,m)} s)}$ time.

There is absolutely nothing new in this proof. The algorithm is described as a series of $s^{O(\log^{f(d,m)} s)}$ time transformations, closely following Allender and Gore [1991] in the appropriate places. We just need to point out that the relevant transformations are still efficiently computable when the algorithm is given an arbitrary input circuit.

PROOF. Let C be the given circuit. Note that by increasing the size of C from s to $s^{O(d)}$, we may always assume at any point in the algorithm that C is a tree (i.e., all gates have fan-out 1).

Transformation 1. Let s be the size of C . We transform C into a probabilistic circuit C' that has $\text{poly}(\log s)$ probabilistic inputs, such that C' has constant depth, $s^{O(1)}$ size, C' has no OR or MOD_m gates for any composite m , and C' has AND gates of fan-in at most $\text{poly}(\log s)$. (The circuit C' is said to accept an input x if it outputs 1 on the majority of settings to the probabilistic inputs.)

First, note that one can replace NOT gates by MOD_m gates (for any m), and one can replace the AND and OR gates by fixed-depth probabilistic circuits with having only MOD_m gates and AND gates of $\text{poly}(\log s)$ fan-in. In fact, all of the AND and OR gates can share the *same* set of $\text{poly}(\log s)$ probabilistic inputs. This is a standard trick that goes back to Valiant and Vazirani [1986] (also found in an alternative proof of Toda's theorem [Kannan et al. 1993]) that can be performed in $s^{O(1)}$ time.

The MOD_m gates for composite m are eliminated as follows. Let $p_1^{e_1} \cdots p_k^{e_k}$ be the factorization of m . Since m divides a number x iff $p_i^{e_i}$ divides x for all i , every MOD_m gate can be replaced by an AND of $\text{MOD}_{p_i^{e_i}}$ gates. (Note the factorization of m can be computed in at most $O(\sqrt{m} \text{poly}(\log m))$ time. Assuming, for example, that $m \leq s$ and m has at most $\text{poly}(\log s)$ distinct primes, this does not factor into the final running time, and does not create ANDs of large fan-in.) Observe that p^e divides a number x iff for all $i = 0, \dots, e-1$, p divides $\binom{x}{p^i}$. Using this fact, a MOD_{p^e} gate can be replaced with a constant-fan-in AND of MOD_p gates of constant-fan-in ANDs, as follows. A MOD_{p^e} gate with t inputs is replaced with an AND of fan-in e , where the inputs are MOD_p gates. For all $i = 0, \dots, e-1$, the i th MOD_p gate has fan-in $\binom{t}{p^i}$, one for every subset of the t inputs that has cardinality p^i . For all $j = 1, \dots, \binom{t}{p^i}$, the i th MOD_p gate has its j th input connected to an AND of the p^i -subset of t inputs corresponding to integer j . All of this can be computed within $t^{O(p^e)}$ time, and hence $s^{O(1)}$ time.

Transformation 2. We have a probabilistic circuit C' with $\text{poly}(\log s)$ probabilistic inputs, constant depth, $s^{O(1)}$ size, no OR or MOD_m gates for any composite m , and AND gates of fan-in at most $\text{poly}(\log s)$. Now we produce a C'' with no probabilistic inputs and all of these properties except that the output gate is now a MAJORITY gate (which outputs the majority value of its inputs). This is easy to do, by enumerating through all possible values of the $\text{poly}(\log s)$ inputs, making a new copy of C' for every valuation, and taking the MAJORITY of all these copies. Certainly the new circuit C'' has size $s^{O(\text{poly}(\log s))}$ and the transformation can be performed in this much time.

Transformation 3. Now we have a circuit C'' has size $s^{O(\text{poly}(\log s))}$ size, a MAJORITY gate at the output, no OR or MOD_m gates for any composite m , and AND gates of fan-in at most $\text{poly}(\log s)$. We produce another constant-depth C''' where all these $\text{poly}(\log s)$ fan-in AND gates are at the bottom: no MOD_p gates are below them in C''' .

Take any AND gate g with $f = \text{poly}(\log s)$ fan-in. Without loss of generality, g has MOD_p gates h_1, \dots, h_f as input for some fixed prime p , by inserting “dummy” MOD_p gates in the appropriate places, and all MOD_p gates have the same fan-in $f' \leq s^{O(\text{poly}(\log s))}$, by inserting “dummy” zeroes in the inputs. We want to show that this AND of MOD_p gates can be rewritten as a MOD_p of ANDs.

Let x_{ij} represent the j th input to the MOD_p gate h_i . Allender and Gore [1991] show that this AND of MOD_p can be rewritten as:

$$\bigwedge_{i=1}^f \left[\sum_j x_{ij} \equiv 0 \pmod{p} \right] = \sum_{k=1}^f (p-1)^{k-1} \sum_{\{i_1, \dots, i_k\} \subseteq [f]} \sum_{\substack{\{j_{1,1}, \dots, j_{1,p-1}\} \in [f']^{p-1} \\ \dots \\ \{j_{k,1}, \dots, j_{k,p-1}\} \in [f']^{p-1}}} \prod_{t=1}^k \prod_{\ell=1}^{p-1} x_{i_t, j_{t,\ell}} \pmod{p}, \quad (4)$$

where $\llbracket P \rrbracket = 1$ if the predicate P is true, and 0, otherwise.

The right-hand side can be represented with a MOD_p gate with fan-in at most

$$O \left(\sum_{k=1}^f \binom{f}{k} (f')^{k(p-1)} \right) \leq s^{O(\text{poly}(\log s))},$$

which is connected to ANDs of fan-in at most $f \cdot (p-1)$. The transformation takes $s^{O(\text{poly}(\log s))}$ time.

Transformation 4. We have a C''' of $s^{O(\text{poly}(\log s))}$ size with AND gates of polylog fan-in connected to the inputs, a MAJORITY at the output, and MOD_{p_i} gates in between, where p_i is a prime dividing $m = O(1)$. We now show how to express C''' as a symmetric function of $s^{O(\text{poly}(\log s))}$ AND gates, completing the proof.

To do this, we prove that if you have a circuit D which has a symmetric function at the output, ANDs at the bottom, and depth- d subcircuits of MOD_{p_i} 's in between, then this can be turned into an equivalent D' with quasipolynomial size, a symmetric function at the top, ANDs at the bottom, and depth- $(d-1)$ subcircuits of MOD_{p_i} 's. That is, the topmost layer of MOD_{p_i} 's can be "consumed" by choosing a different symmetric function.

We may assume without loss of generality that all f gates with input to the symmetric function $F : [f] \rightarrow \{0, 1\}$ are MOD_p gates, for a fixed prime p , and all of the MOD_p gates have the same fan-in f' (by adding dummy wires and gates where necessary). Let x_{ij} be the j th input to the i th MOD_p gate. Note that the function we want to simulate is $H(x_{1,1}, \dots, x_{f,f'}) = F(\sum_{i=1}^f \text{MOD}_p(x_{i,1}, \dots, x_{i,f'}))$. We will replace H with a symmetric function F' of ANDs of polylogarithmic fan-in. Then, applying Transformation 3 to these ANDs, the resulting circuit can be converted into one which has the ANDs at the bottom and only a quasipolynomial increase in size.

Define

$$G(x_{1,1}, \dots, x_{f,f'}) := F \left(\sum_{i=1}^f \text{MOD}_p(x_{i,1}, \dots, x_{i,f'}) \bmod p^k \right),$$

where k is the smallest integer exceeding $\log_p f$. Then, $p^k > f$, so it is clear that $G(x_{1,1}, \dots, x_{f,f'}) = H(x_{1,1}, \dots, x_{f,f'})$ when all x_{ij} are in $\{0, 1\}$. We shall show how to implement G as a symmetric function F' of ANDs.

We use the *modulus amplifying polynomials* of Beigel and Tarui [1994]. Define

$$P_k(x) = (-1)^k (x-1)^k \left(\sum_{i=0}^{k-1} \binom{k+i-1}{i} x^i \right) + 1.$$

This polynomial has the property that for all $x \geq 0$ and $p \geq 1$,

$$x = 0 \bmod p \implies P_k(x) = 0 \bmod p^k,$$

$$x = 1 \bmod p \implies P_k(x) = 1 \bmod p^k.$$

Defining $Q_k(x) = 1 - P_k(x^{p-1})$ and appealing to Fermat's little theorem, it follows that $Q_k(x) = 1 \bmod p^k$ if p divides x , and is equal to $0 \bmod p^k$ otherwise. Therefore, $Q_k(\sum_{i=1}^{f'} y_i) = \text{MOD}_p(y_1, \dots, y_{f'}) \bmod p^k$, and

$$G(x_{1,1}, \dots, x_{f,f'}) = F \left(\sum_{i=1}^f Q_k \left(\sum_{j=1}^{f'} x_{i,j} \right) \bmod p^k \right).$$

Note that each $Q_k(\sum_{j=1}^{f'} x_{i,j})$ is a symmetric multivariate polynomial of degree at most $k(p-1)$. Hence, Q_k can be expanded into a sum of at most $(f \cdot f')^{O(k(p-1))} \leq s^{O(\text{poly}(\log s))}$ terms. Each term is a product of $\text{poly}(\log s)$ variables and a coefficient c that is represented in $O(k \log k) \leq \text{poly}(\log s)$ bits and easily computed. The product of variables can be directly represented by an AND. Multiplication by the coefficient c can be simulated by taking the sum of c copies of the relevant monomials (ANDs).

Therefore, the sum of all f of these sums of monomials can be efficiently expressed as a single sum modulo p^k of $s^{O(\text{poly}(\log s))}$ AND gates, where each AND has fan-in $k(p-1) \leq \text{poly}(\log s)$. Finally, we take the symmetric function F' to be: compute the sum v of the outputs of all the AND gates created, then output $F(v \bmod p^k)$. Observe that a symmetric function composed with a sum modulo p^k is still a symmetric function.

In summary, for any constant depth circuit, all of these transformations take at most quasipolynomial time, increase the circuit size by only a quasipolynomial amount, and the transformations are applied at most a quasipolynomial number of times. (Transformation 4 is applied a constant number of times.) Moreover, the symmetric function generated at the end of the process takes no more time to evaluate than the time it takes to build the SYM^+ circuit. In more detail, the final symmetric function has the form

$$F(v) = \text{MAJORITY}((\dots((v \bmod p_1^{k_1}) \bmod p_2^{k_2}) \dots \bmod p_{d'}^{k_{d'}}),$$

for some d' that depends on the constant depth d and constant modulus m . Here, MAJORITY outputs the high-order bit of its input, and each $p_i^{k_i}$ is at most a constant factor larger than the size of the final circuit. \square

B. TRANSFORMING TO SYM^+ WITHOUT NEGATIONS

Suppose we are given a circuit C which is a SYM of $K = n^{\text{poly}(\log s)}$ ANDs of $\text{poly}(\log s)$ fan-in, where the AND gates may have some negated variables as input. Here, we sketch how to transform C into a circuit C' which is a SYM of $K' = n^{\text{poly}(\log s)}$ ANDs of $\text{poly}(\log s)$ fan-in, where the ANDs are of variables only (no negations).

The SYM of ANDs circuit C can be represented as a function $g : \{0, \dots, K\} \rightarrow \{0, 1\}$ composed with a $\text{poly}(\log n)$ -degree polynomial $h : \{0, 1\}^n \rightarrow \{0, 1\}$. The function g simulates the SYM gate, and the function h sums up the total number of ANDs which are true on a given assignment. That is, h is a sum of terms over all possible AND gates, where each term directly corresponds to the product of the literals in some AND gate: each negated literal $\neg x_i$ is multiplied as $(1 - x_i)$, and each positive literal x_i is multiplied as just x_i . Hence, a term evaluates to 1 on a 0 – 1 assignment if and only if the corresponding AND outputs true.

We can expand the polynomial h into a sum of products of variables in $O(n^{\text{poly}(\log n)})$ time, since h is a sum of $n^{\text{poly}(\log n)}$ products, each product being a $\text{poly}(\log n)$ -degree polynomial. Call this new arithmetic expression h' , which has the form

$$h'(x_1, \dots, x_n) = \sum_{S \subseteq [n]} c_S \cdot \prod_{i \in S} x_i,$$

where each c_S is an integer in the interval $[-K', K']$, and all but $O(K')$ of the c_S are zero (and $K' = n^{\text{poly}(\log n)}$). Say that a subset $S \subseteq [n]$ is *negative* if $c_S < 0$, and *positive* if $c_S > 0$.

It is easy to represent h' as a SYM of ANDs of *entirely nonnegated variables* (with a different SYM function from before). The trick is to expand the domain of the function g mildly so that it can keep track of the contributions from negative S and positive S , separately. Let ℓ be the smallest integer satisfying $2^\ell > K'$. Initially, our new circuit C' has no AND gates. For every S which is negative, add $-c_S$ AND gates over the variables in S to the circuit C' . For every S which is positive, add $2^\ell \cdot c_S$ AND gates over the variables in S to the circuit C' . Now consider the sum of all AND gates in C' on some assignment $(y_1, \dots, y_n) \in \{0, 1\}^n$. This sum has the form

$$AB$$

in binary, where B is an ℓ -bit string and $A - B$ (construing A and B as nonnegative integers) equals $h(y_1, \dots, y_n)$. Hence, we can redefine the function g to be

$$g'(AB) = A - B.$$

This g , combined with this description of AND gates, represents a SYM of ANDs of non-negated variables that is equivalent to the original circuit C .

C. COPPERSMITH'S ALGORITHM

Recall we are studying the following algorithm of Coppersmith.

LEMMA 6.1 [COPPERSMITH 1982]. *For all sufficiently large N , multiplication of an $N \times N^{-1}$ matrix with an $N^{-1} \times N$ matrix can be done in $O(N^2 \log^2 N)$ arithmetic operations.*

Prima facie, it could be that Coppersmith's algorithm is nonuniform, making it difficult to apply. For the sake of completeness, here we verify using standard ideas that Coppersmith's algorithm can indeed be implemented to run (even on a multitape TM) in $O(N^2 \cdot \text{poly}(\log N))$ time, on matrices over any field of $\text{poly}(N)$ elements. (As we work with 0-1 matrices A'' and B'' in our application, it suffices for us to work over a prime field of $\text{poly}(N)$ elements.) We focus on the implementation details of his algorithm, without going very far into its correctness. The algorithm relies on some of the older tools from the matrix multiplication literature. More background on these tools can be found in the highly readable reference [Pan 1984].

Coppersmith's algorithm follows a paradigm introduced by Schönhage [1981]. For example, suppose we wish to multiply two matrices A'' and B'' . First, we *preprocess* A'' and B'' in some efficient way; in our first example, we devise highly structured matrices A, A', B, B' so that $A'' \cdot B'' = A' \cdot A \cdot B \cdot B'$. The matrices A and B are sparse "partial" matrices with particular structure in their nonzeros, and A' and B' are explicit matrices of scalar constants which are independent of A'' and B'' . Next, we recursively apply a constant-sized matrix multiplication algorithm to multiply A and B essentially optimally. (Recall that Strassen's algorithm has an analogous form; such algorithms are known to be efficiently implementable on a multitape TM.) Finally, we *postprocess* the resulting product C to obtain our desired product $A'' \cdot B''$; in the first example, this means computing $A' \cdot C \cdot B'$. Using the explicit structure of A' and B' , these matrix products are also done nearly optimally. Our aim is to verify that each step of this process can be efficiently computed, for Coppersmith's full matrix multiplication algorithm.

Coppersmith begins with A'' of dimensions $2^{4M/5} \times \binom{M}{4M/5} 2^{4M/5}$ and B'' of dimensions $\binom{M}{4M/5} 2^{4M/5} \times 2^{M/5}$ where $M \approx \log N$, and obtains an $O(5^M \text{poly}(M))$ algorithm for their multiplication. Later, he symmetrizes the construction to get the algorithm for the desired dimensions. In this first construction, the structured matrices A' and B' have dimensions $2^{4M/5} \times 2^M$ and $2^M \times 2^{M/5}$, respectively. Coppersmith needs that all $2^{4M/5} \times 2^{4M/5}$ submatrices of A' and $2^{M/5} \times 2^{M/5}$ submatrices of B' are nonsingular. Following Schönhage, this can be accomplished by picking A' and B' to be rectangular Vandermonde matrices. More precisely, the i, j entry of A' is $(\alpha_j)^{i-1}$, where $\alpha_1, \alpha_2, \dots$ are distinct elements of the field; B' is defined analogously. Such matrices have the additional advantages that they can be succinctly described (with 2^M field elements), and linear algebra with them can be done very efficiently, as described here.

The matrices A and B have dimensions $2^M \times 3^M$ and $3^M \times 2^M$, respectively. Although these dimensions are large, the matrices are stored in a sparse representation, and they have structure in their nonzeros. In more detail, A has only $O(5^M)$ nonzeros, B has only $O(4^M)$ nonzeros, and there is an optimal algorithm for multiplying

2×3 (with 5 nonzeros) and 3×2 matrices (with 4 nonzeros) that can be recursively applied to multiply A and B optimally, in $O(5^M \cdot \text{poly}(M))$ operations. (In particular, the 2×3 and 3×2 matrix multiplication is an “approximate” algorithm, which can be recursively applied to larger matrices using $O(M)$ -degree univariate polynomials over the field; operations on such polynomials increase the overall time by only a $\text{poly}(M)$ factor.) These A and B are constructed by multiplying each of the $\binom{M}{4M/5} 2^{4M/5}$ columns in A'' and $\binom{M}{4M/5} 2^{4M/5}$ rows in B'' by inverses of Vandermonde matrices and their transposes (the inverses of appropriate $2^{4M/5} \times 2^{4M/5}$ submatrices of A' and $2^{M/5} \times 2^{M/5}$ submatrices of B' , respectively). Due to the structure of inverse Vandermonde matrices and their transposes, $n \times n$ matrices of this form can be multiplied with n -vectors in $O(n \cdot \text{poly}(\log n))$ operations with explicit algorithms (for references, cf. Canny et al. [1989] and Bini and Pan [1994]).⁶ Hence, the inverse of a submatrix of A' can be multiplied with an arbitrary vector in $O(2^{4M/5} \cdot \text{poly}(M))$ operations. It follows that constructing A and B takes only $O(\binom{M}{4M/5} 2^{4M/5} \cdot 2^{4M/5} \cdot \text{poly}(M))$ time. Since $5^M \approx \binom{M}{4M/5} 4^{4M/5}$ (within $\text{poly}(M)$ factors), this quantity is $O(5^M \cdot \text{poly}(M))$.

By construction (using an efficient correspondence between columns of A'' and columns of A' with $2^{4M/5}$ nonzeros), we have $A'' \cdot B'' = A' \cdot (A \cdot B) \cdot B'$. After A and B are constructed, the constant-sized algorithm for 2×3 and 3×2 mentioned previously can be applied in the usual recursive way to multiply the sparse A and B in $O(5^M \cdot \text{poly}(M))$ time; call this matrix Z . Then, using the Vandermonde structure of A' and B' , the product $Z' = A' \cdot Z$ can be done in $o(5^M \cdot \text{poly}(M))$ operations, and the final product $Z' \cdot B'$ can be done in $o(5^M \cdot \text{poly}(M))$ operations. All in all, we have an algorithm for multiplying matrices of dimensions $2^{4M/5} \times \binom{M}{4M/5} 2^{4M/5}$ and $\binom{M}{4M/5} 2^{4M/5} \times 2^{M/5}$ that is explicit and uses $O(5^M \cdot \text{poly}(M))$ operations. Call this Algorithm 1. Observe Algorithm 1 also works when the entries of A'' and B'' are themselves matrices over the field. (The running time will surely increase in proportion to the sizes of the underlying matrices, but the bound on the number of operations on the entries remains the same.)

We can extract more algorithms from this construction by exploiting the symmetries of bilinear algorithms. The underlying 2×3 and 3×2 matrix multiplication algorithm with five products is a bilinear algorithm, meaning that it can be expressed in the so-called trilinear form

$$\sum_{ijk} A_{ik} B_{kj} C_{ji} + p(x) = \sum_{\ell=1}^5 \left(\sum_{ij} \alpha_{ij} A_{ij} \right) \cdot \left(\sum_{ij} \beta_{ij} B_{ij} \right) \cdot \left(\sum_{ij} \gamma_{ij} C_{ij} \right), \quad (5)$$

where α_{ij} , β_{ij} , and γ_{ij} are constant-degree polynomials in x over the field, and $p(x)$ is a polynomial with constant coefficient 0. Such an algorithm can be converted into one with no polynomials and minimal extra overhead (as described in Coppersmith’s paper). Typically, one thinks of A_{ik} and B_{kj} as entries in the input matrices, and C_{ji} as indeterminates, so the left-hand side of (5) corresponds to a polynomial whose C_{ji} coefficient is the ij entry of the matrix product. Note the *transpose* of the third matrix C corresponds to the final matrix product. The right-hand side corresponds to the special

⁶In general, operations on Vandermonde matrices, their transposes, their inverses, and the transposes of inverses can be reduced to fast multipoint computations on univariate polynomials. For example, multiplying an $n \times n$ Vandermonde matrix with a vector is equivalent to evaluating a polynomial (with coefficients given by the vector) on the n elements that comprise the Vandermonde matrix, which takes $O(n \log n)$ operations. This translates to $O(n \cdot \text{poly}(\log n))$ time on multitape TMs over small fields.

matrix multiplication algorithm with only 5 products. For example, Strassen's famous 7-multiplication algorithm can be expressed in the form of (5) as follows:

$$\begin{aligned} \sum_{i,j,k=0,1} A_{ik} B_{kj} C_{ji} = & (A_{00} + A_{11})(B_{00} + B_{11})(C_{00} + C_{11}) \\ & + (A_{10} + A_{11})B_{00}(C_{01} - C_{11}) + A_{00}(B_{01} - B_{11})(C_{10} + C_{11}) \\ & + (A_{10} - A_{00})(B_{00} + B_{01})C_{11} + (A_{00} + A_{01})B_{11}(C_{10} - C_{00}) \\ & + A_{11}(B_{10} - B_{00})(C_{00} + C_{01}) + (A_{01} - A_{11})(B_{10} + B_{11})C_{00}. \end{aligned} \quad (6)$$

The left-hand side of (5) and (6) represents the trace of the product of three matrices A , B , and C (where the ij entry of matrix X is X_{ij}). It is well known that every bilinear algorithm naturally expresses multiple algorithms through this trace representation. Since

$$\text{tr}(ABC) = \text{tr}(BCA) = \text{tr}(CAB) = \text{tr}((ABC)^T) = \text{tr}((BCA)^T) = \text{tr}((CAB)^T),$$

if we think of A as a symbolic matrix and consider (5), we obtain a new algorithm for computing a matrix A when given B and C . Similarly, we get an algorithm for computing a B when given A and C , and analogous statements hold for computing A^T , B^T , and C^T . So the aforementioned algorithm for multiplying a sparse 2×3 and sparse 3×2 yields several other algorithms. In particular (the case of computing B^T from A^T and C), we obtain an algorithm for computing four entries in a 3×2 matrix which is the product of a 3×2 matrix (with five nonzeros) and a 2×2 matrix.

Using the identity $\text{tr}(ABC) = \text{tr}((BCA)^T) = \text{tr}(A^T C^T B^T)$, we can treat B^T as symbolic and let A^T and C^T correspond to input matrices in (5). Applying the resulting algorithm recursively, a very similar preprocessing and postprocessing can be used to multiply $\binom{M}{4M/5} 2^{4M/5} \times 2^{4M/5}$ and $2^{4M/5} \times 2^{M/5}$ matrices using an algorithm that runs in $O(5^M \cdot \text{poly}(M))$ time over a small field.

In more detail, recall in Algorithm 1 the matrices A'' and B'' were decomposed to satisfy $A'' \cdot B'' = A' \cdot A \cdot B \cdot B'$. The trace identity tells us

$$\text{tr}(A'' B'' \cdot C) = \text{tr}(A' A \cdot B B' \cdot C) = \text{tr}(B \cdot B' C A' \cdot A) = \text{tr}(A^T \cdot (A')^T C^T (B')^T \cdot B^T).$$

This suggests the following algorithm for multiplying $\binom{M}{4M/5} 2^{4M/5} \times 2^{4M/5}$ and $2^{4M/5} \times 2^{M/5}$ matrices. Given A'' and C'' of the appropriate dimensions, preprocess C'' into the $2^M \times 2^M$ matrix $D = B' \cdot (C'')^T \cdot A'$, and use A' as before to preprocess A'' into a sparse $3^M \times 2^M$ matrix A^T having $\binom{M}{4M/5} 4^{4M/5} \approx 5^M$ nonzeros. Both steps can be done efficiently using the Vandermonde structure of A' and B' . Next, multiply A^T and D^T , following the bilinear algorithm for computing a 3×2 (with four nonzeros) from a 3×2 (with five nonzeros) and a 2×2 , in $O(5^M \text{poly}(M))$ time. The multiplication results in a $3^M \times 2^M$ matrix B with $O(4^M)$ nonzeros, which can be efficiently transformed to the output matrix using the inverses of submatrices of B' . (This postprocessing step is analogous to the preprocessing of B'' in Algorithm 1.) Notice we have analogous preprocessing, multiplication, and postprocessing steps, albeit the steps are “out of order” from before. (Before, multiplication of the result matrix C by A' and B' occurred in postprocessing; now it occurs in preprocessing, as C is now part of the input.) Call this construction Algorithm 2.

Next, we may “tensorize” the two algorithms in a standard way. This consists of dividing the input matrices into blocks, executing Algorithm 1 on the blocks themselves, and calling Algorithm 2 when the product of two blocks is needed. As both of these algorithms are explicit and efficient, their “tensorization” is also explicit and efficient.

Algorithm 1 multiplies $2^{4M/5} \times \binom{M}{4M/5} 2^{4M/5}$ and $\binom{M}{4M/5} 2^{4M/5} \times 2^{M/5}$ matrices, and Algorithm 2 multiplies $\binom{M}{4M/5} 2^{4M/5} \times 2^{4M/5}$ and $2^{4M/5} \times 2^{M/5}$. Hence, their tensorization multiplies matrices of dimensions $(2^{4M/5} \cdot \binom{M}{4M/5} 2^{4M/5}) \times (\binom{M}{4M/5} 2^{4M/5} \cdot 2^{4M/5})$ and $(\binom{M}{4M/5} 2^{4M/5} \cdot 2^{4M/5}) \times (2^{M/5} \cdot 2^{M/5})$, and the algorithm runs in $O(5^{2M} \cdot \text{poly}(M))$ time. Since $\binom{M}{4M/5} 4^{4M/5} \approx 5^M$, this means we are multiplying $5^M \times 5^M$ and $5^M \times 2^{2M/5}$ in $O(5^{2M} \text{poly}(M))$ time. Call this Algorithm 3. This is the algorithm obtained by Coppersmith.

Finally, using the symmetry of Algorithm 3 itself, we can obtain an algorithm for multiplying a $5^M \times 2^{2M/5}$ matrix with a $2^{2M/5} \times 5^M$ matrix in $O(5^{2M} \text{poly}(M))$ time. Algorithm 3 is also a bilinear algorithm that can be interpreted as an efficient way to compute $\text{tr}(ABC)$ where A is $5^M \times 5^M$, B is $5^M \times 2^{2M/5}$, and C is $2^{2M/5} \times 5^M$. In this version of Algorithm 3, we have treated A and B as input, and C as symbolic. Treating B and C as input yields an algorithm for multiplying $5^M \times 2^{2M/5}$ and $2^{2M/5} \times 5^M$ in $O(5^{2M} \text{poly}(M))$ time. This algorithm also has a preprocessing step, a product of partial matrices, then a postprocessing step, which involve multiplications with Vandermonde-style matrices, their transposes, their inverses, and their inverse transposes. The important point is that this transformation does not fundamentally change the algorithm: just as Algorithm 2 is a “reordering” of Algorithm 1, this transformation of Algorithm 3 only reorganizes these efficiently computable operations. It follows the final algorithm will also be efficiently computable. (Of course, it is possible in principle to describe this algorithm directly as a preprocessing-multiplication-postprocessing procedure, but it is quite messy.) Let $N = 5^M$. We have arrived at the following.

COROLLARY C.2. *For all sufficiently large N , two 0-1 matrices of dimensions $N \times N^{-1}$ and $N^{-1} \times N$ can be multiplied over the integers in $O(N^2 \cdot \text{poly}(\log N))$ time.*

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