Diagonalization in proof complexity

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Abstract

We study diagonalization in the context of implicit proofs of [10]. We prove that at least one of the following three conjectures is true:

- There is a function $f : \{0,1\}^* \to \{0,1\}$ computable in $\mathcal{E}$ that has circuit complexity $2^{\Omega(n)}$.
- $\mathcal{NP} \neq co\mathcal{NP}$.
- There is no p-optimal propositional proof system.

We note that a variant of the statement (either $\mathcal{NP} \neq co\mathcal{NP}$ or $\mathcal{NE} \cap co\mathcal{NE}$ contains a function $2^{\Omega(n)}$ hard on average) seems to have a bearing on the existence of good proof complexity generators. In particular, we prove that if a minor variant of a recent conjecture of Razborov [17, Conjecture 2] is true (stating conditional lower bounds for the Extended Frege proof system EF) then actually unconditional lower bounds would follow for EF.

The only method for demonstrating unprovability of a Π^0_1-sentence in a theory (containing some amount of arithmetic) is diagonalization. One would like to adapt this to a non-uniform setting in which theories are replaced by propositional proof systems and Π^0_1-sentences by tautologies.

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This fails, at least in the straightforward adaptation. In particular, many strong proof systems do prove their own consistency by polynomial size proofs (cf.[5, Chpt.14]). However, there is a non-uniform setting in first-order theories where diagonalization gives non-trivial lower bounds.

To explain this we need to fix some notation first. By a theory we shall mean a set of axioms, not the set of its consequences. We shall assume (w.l.o.g.) that the language of all theories in question is the language of bounded arithmetic theory S^1_2. For a natural number N the symbol $\underline{N}$ is the dyadic numeral inductively defined by: $\underline{0} := 0$, $\underline{1} := 1$, $\underline{2} := (1+1)$, $\underline{2k} := (\underline{2} \cdot \underline{k})$, and $\underline{2k+1} := (\underline{2k} + 1)$.

Given a theory T definable in the language of S^1_2, let $Prf_T(y,x)$ be a formula in the language of S^1_2 expressing that y is a T-proof of formula x. There is a canonical formalization of syntax of logic (terms, formulas, proofs, etc.) and of the notion of provability (see [13] or [5] for details), and its usual properties are provable in S^1_2. Assuming that $T \supseteq S^1_2$, as we shall do, thus allows us to use this formalization in T too.

Define a diagonal formula $A(x)$ satisfying:

$$A(x) \ \equiv \ \forall y, |y| \leq s(x) \longrightarrow \neg Prf_T(y, \ulcorner A(\dot{x}) \urcorner).$$

where $s(x)$ is any term and $\ulcorner A \urcorner$ is a (canonical) number encoding A such that its length is proportional to the size of A, and $\dot{x}$ is the formalization of dyadic numerals. Then, by the standard argument, we have the following theorem. (The condition that $T \in \mathcal{NP}$ is technical and implies, in particular, that T is definable and that the binary relation $Prf_T(y,x)$ is in $\mathcal{NP}$ and definable by a Σ^b_1-formula, and that all its true instances have polynomial size proofs in S^1_2, cf.[5, p.303].)

Theorem 0.1 *Let $T \supseteq S^1_2$ be a consistent theory such that $T \in \mathcal{NP}$. Then for any $N \geq 1$ the sentence $A(\underline{N})$ is true and provable in T but any T-proof of the sentence must have the size at least $s(N)$.*

Note that the length of the formula $A(\underline{N})$ is $O(\log N)$ and hence the lower bound is non-trivial, as long as $s(N) >> \log(N)$. A foremost example of this reasoning is the Finitistic Gödel's theorem of Friedman [3] and Pudlák [13, 14] (the theorem says that any T-proof of a formula $Con_T(\underline{N}) := \forall y, |y| \leq \underline{N} \rightarrow \neg Prf(y, \ulcorner 0 = 1 \urcorner)$ expressing the consistency of T w.r.t. proofs of length at most N, must have the size at least $N^{\Omega(1)}$). Other, quantitatively more subtle applications, can be found in bounded arithmetic, cf.[5, Chpt.10].

One would like to adapt this to propositional proof complexity. The problem with this is not that we deal with first-order theories (as we know how to pass between them and propositional proof systems) but with the fact that $A(x)$ is not a bounded formula and the instances $A(\underline{N})$ translate into propositional formulas of length at least $s(N)$, which is super-polynomial in order to get non-trivial lower bound. An idea suggests itself at this point: Try to use the concept of implicit proofs from [10]. This is what we investigate in this paper. A variant of the idea of implicit proofs is recalled in Section 1. In Section 2 we derive the theorem mentioned in the abstract. Its variant is derived in Section 3 and linked to proof complexity generators in Section 4.

The paper is self-contained, assuming that the reader has a general background in bounded arithmetic and proof complexity (only basic things are assumed), see [5]. Let us just recall the terminology and some more notation. $Time(t(n))$ is the class of languages computable in deterministic time $O(t(n))$, $\Sigma_i Time(t(n))$ is the Σ_i-level of the time $O(t(n))$ hierarchy, $NTime(t(n)) = \Sigma_1 Time(t(n))$, and $\mathcal{E} = Time(2^{O(n)})$. For a function $f : \{0,1\}^* \to \{0,1\}$, $C_f(n)$ is the circuit complexity of computing f on $\{0,1\}^n$, while $H_f(n)$ is it's hardness on average in the sense of [12]. A "proof system" tacitly means a "propositional proof system": It is a non-deterministic acceptor of the set of propositional tautologies in the DeMorgan language. A proof of a tautology in a proof system is any particular computation of the proof system accepting the formula. A proof system is p-bounded iff the proof system accepts all tautologies in a fixed polynomial time, and it is p-optimal iff it polynomially simulates other proof systems, cf.[2]. The length of a string w is denoted $|w|$, and a number is identified with the string of its bits (i.e. $|m| \sim \log(m)$).

A prominent proof system is the Extended Frege proof system EF. It is the usual Hilbert style calculus based on a finite number of axiom schemes and inference rules (sound and implicationally complete in a complete language) that has an additional ability (via the so called Extension Rule) to abbreviate large formulas occurring in a proof by new atoms, see [2, 5] for details. Alternatively one can view EF as a Hilbert style calculus operating with circuits rather than only with formulas. The prominence of EF comes (at least in our context) from its relation to theory S^1_2.

Finally let us recall a well-known translation of formulas into propositional formulas, just for the case of $A(x)$. Given $N \geq 1$, there is a propositional 3DNF formula denoted $||A(x)||_N$ of size $s(N)^{O(1)}$ that expresses (by the fact of being a tautology) that $A(N)$ is true. It is constructed as in the proof of the $\mathcal{NP}$-completeness of satisfiability as follows. Fix a particular Σ_1^b-formula ($\mathcal{NP}$-definition) Prf_T. Formula $||A(x)||_N$ has $s(N)$ atoms for bits of a potential y, $s(N)^{O(1)}$ atoms for bits of a potential witness z to the validity of the $\mathcal{NP}$-statement $Prf_T(y, \lceil A(\underline{N}) \rceil)$, and additional $s(N)^{O(1)}$ auxiliary atoms. The auxiliary atoms are used for naming bits in the canonical computation of the truth value of $Prf_T(y, \lceil A(\underline{N}) \rceil)$ with the witness z. Formula $||A(x)||_N$ is built by induction on the logical complexity of A and says that if all local conditions in the computation are satisfied then $Prf_T(y, \lceil A(\underline{N}) \rceil)$ fails. See [5, Chpt.9] for details of the translation.

1 Implicit proofs of implicit formulas

The idea of implicit proofs is to represent a proof in a proof system, a binary string, not by the string itself but by a circuit that computes any individual bit of the proof knowing only its position but not the whole proof. A proof can be, in principle, exponentially larger than a circuit representing it. Hence a proof system operating with such circuits instead of directly with proofs can be quite powerful. This concept has been proposed and studied in [10]. In this paper we borrow the general idea of representing long proofs (and, in fact, also long formulas) by small circuits but otherwise we do not use anything from [10]. Let us now proceed formally.

Let w be a 0-1 string of length 2^k. Identify $i < 2^k$ with vectors $i = (i_1, \ldots, i_k) \in \{0,1\}^k$ ordered lexicographically. We say that a circuit $C(x_1, \ldots, x_k)$ represents w iff $C(i_1, \ldots, i_k) = w_i$ (the ith bit of w), for all $i < 2^k$. Similarly, if W is a 0-1 $2^k \times 2^k$ matrix then we say that a circuit $D(x_1, \ldots, x_k, y_1, \ldots, y_k)$ represents W iff $D(i,j) = W_{i,j}$, for all $i, j < 2^k$.

Let M be a non-deterministic polynomial time machine. For any input w of length 2^ℓ and represented by a circuit $C(z_1, \ldots, z_\ell)$, for suitable $k = O(\ell)$ (given by the time of M) and any $2^k \times 2^k$ matrix W represented by a circuit $D(x_1, \ldots, x_k, y_1, \ldots, y_k)$ we can write down a propositional formula $\sigma^M_{C,D}$ expressing that W is an accepting computation of M (given by W listing in its rows all instantaneous descriptions of the computation) on input w. Moreover, the size of $\sigma^M_{C,D}$ is only proportional to the sizes of C and D. Such a formula can be constructed as follows.

The formula has atoms $x_1, \ldots, x_k, y_1, \ldots, y_k, z_1, \ldots, z_\ell$ for inputs of D and C respectively. Using these atoms any local condition imposed on W

by M (there is a finite number of them) can be written by a circuit of size proportional to the sizes of C and D (as a local condition speaks only about a constant number of positions in W). This means that the circuit computes constantly value 1 (for all evaluations of the atoms, i.e. for all positions in W) iff all local conditions are satisfied on the whole of W. Using additional auxiliary atoms for values of subcircuits of C and D this can be expressed by a formula (as in the proof of the $\mathcal{NP}$-completeness of satisfiability). To summarize: $\sigma^M_{C,D}$ is a tautology iff W is indeed an accepting computation of M on w.

We will need one additional formula, this time first-order. Assume that a circuit $C(x_1, \ldots, x_k)$ represents a 3DNF formula φ_C (of size at most 2^k). Let $\lceil C \rceil$ be its number code. Then there is a formula $BigTaut(x)$ such that $BigTaut(\lceil C \rceil)$ formalizes that φ_C is a tautology. This can be expressed without going through φ_C: For any truth assignment y (of the length bounded by a suitable term $|C|^{O(1)}$) there is a term in the 3DNF formula represented by C (this quantifies over inputs of C and not over the whole big formula) satisfied by y.

We formulate the following lemma for the diagonal formula $A(x)$ only, but it holds for any formula of a similar logical form.

Lemma 1.1 *There is a polynomial time function $g(x)$ such that, for any $N \geq 1$, $g(N)$ is (the number code of) a circuit representing (in the sense as above) formula $||A(x)||_N$.*

Moreover, the valid implication

$$BigTaut(g(x)) \longrightarrow A(x)$$

is provable in theory S^1_2.

Proof :

The existence of such a function g follows from the high uniformity (for a fixed formula $A(x)$) of the construction of $||A(x)||_N$ as recalled at the end of the introductory section. In fact, there is a polynomial time function $g^*_A(x, y)$ such that $g^*_A(N, i)$ computes the ith bit of $||A(x)||_N$ ($g(N)$ is then the circuit computing $g^*_A(N, y)$ for y bounded by a suitable term $N^{O(1)}$). Function g^*_A is defined by induction on the logical complexity of A from analogous functions g^*_B for subformulas B of A. The claim for the base case of atomic formulas follows as the local conditions are the same across the whole computation (of the values of terms in the atomic formula) and the algorithm computing

g^*_B just needs to know a position in the computation tableaux (i.e. the index i) in order to use the appropriate atoms while writing down the particular local conditions in the propositional translation.

The validity of the implication in the second part of the lemma is clear from the definition of the propositional translation. To see that it is also provable in S^1_2 reason as follows. Assume that y is a witness for the failure of $A(x)$: $|y| \leq x \wedge Prf_T(y, \ulcorner A(\dot{x}) \urcorner)$. Then there is a truth assignment $f(x,y)$, computed from x and y by a polynomial time function f, to atoms of $||A(x)||_x$ that violates the formula (cf. [5, Sec.9.3] for an analogous statement and a detailed discussion). We claim that S^1_2 proves (from the assumption that y witnesses the failure of $A(x)$) that "$f(x,y)$ does not satisfy any term in the 3DNF formula represented by $g(x)$" (and hence $\neg BigTaut(g(x))$). This is showed again by induction on logical complexity of A, the base case being that of atomic formulas. For them the claim is proved by the induction on the size of circuits computing the terms in the atomic formula, i.e. by the length induction on parameter y.

q.e.d.

Note that the circuit $g(N)$ has size, by the fact that g is polynomial time computable, at most $\log(N)^{O(1)}$.

2 A theorem

Theorem 2.1 *At least one of the following three statements is true:*

(i) There is a function $f : \{0,1\}^* \rightarrow \{0,1\}$ computable in $\mathcal{E}$ that has circuit complexity $2^{\Omega(n)}$.

(ii) $\mathcal{NP} \neq co\mathcal{NP}$.

(iii) There is no p-optimal propositional proof system.

Proof :

Before we start with the actual proof let us recall a well-known technical fact (see e.g. [5, p.303]) that will be used several times in the proof (in (2), (9) and (11)): Any true Σ^b_1-sentence has a polynomial size proof in S^1_2. This is proved by induction on the logical complexity of the sentence.

Let us begin with the proof proper. We shall assume that all three statements are false and derive a contradiction via Theorem 0.1.

(1) Let P be a proof system witnessing that both (ii) and (iii) fail: P is p-bounded and also p-optimal. Assume w.l.o.g. that P contains EF. Let $Prf_P(u,v)$ be a Σ_1^b-formula formalizing that "u is a P-proof of formula v".

(2) Define theory T to be S_2^1 augmented by an extra axiom, a form of reflection principle:

$$\forall x, y, u,\ Prf_P(u, \ulcorner \sigma_{x,y}^P \urcorner) \longrightarrow BigTaut(x)\ .$$

Note that the antecedent of the implication is a Σ_1^b-formula because the (code of) formula $\sigma_{x,y}^P$ is polynomial time computable form x, y (for fixed P), cf. Section 1. This implies, in particular, that all true instances of the formula have polynomial size proofs in S_2^1.

(3) Let $A(x)$ be the diagonal formula from the introduction, with the term $s(x)$ being simply x. Consider the propositional formula $||A(x)||_N$. The formula is a tautology as $A(N)$ is true. The size of the formula is $N^{O(1)}$ but by Lemma 1.1 there is a circuit $C_N := g(N)$ of size $n^{O(1)}$, $n := \log(N)$, representing the formula (in the sense of Section 1).

(4) The set of all formulas $||A(x)||_N$, $N \geq 1$, is polynomial time decidable and hence we can use it as axioms in some proof system. By the hypothesis that P is p-optimal there is a (deterministic) polynomial time algorithm M computing from the string $||A(x)||_N$ a P-proof of $||A(x)||_N$.

(5) The output of M is a particular accepting computation of P, i.e. an $2^{O(n)} \times 2^{O(n)}$ matrix W^N encoding the computation. As M runs in deterministic polynomial time, $W_{i,j}^N$ as a function of $i, j \in \{0,1\}^{O(n)}$ is in $\mathcal{E}$.

(6) Assuming that also statement (i) fails, there exists a circuit $D(i,j)$ in 2-times $O(n)$ variables and of size $2^{\delta \cdot n}$ that represents W^N, for arbitrary small $\delta > 0$. We shall choose a particular δ in the step (12).

(7) Take an instance of the reflection principle by substituting for x and y the codes of D and C_N respectively:

$$\forall u,\ Prf_P(u, \ulcorner \sigma_{\ulcorner C_N \urcorner, \ulcorner D \urcorner}^P \urcorner) \longrightarrow BigTaut(\ulcorner C_N \urcorner)\ .$$

(8) By Section 1 the size of $\sigma_{C_N,D}^P$ is polynomial in the sizes of C_N and D, i.e. it is $2^{O(\delta \cdot n)}$. Now we use the hypothesis that P is also p-bounded. Hence there is a P-proof e of $\sigma_{C_N,D}^P$ of size $2^{O(\delta \cdot n)}$. Note that the constants implicit in the O-notation are fixed and independent of δ. Substituting $\lceil e \rceil$ for u in the formula in (7) we get:

$$Prf_P(\lceil e \rceil, \lceil \sigma^P_{\lceil C_N \rceil, \lceil D \rceil} \rceil) \longrightarrow BigTaut(\lceil C_N \rceil) \ .$$

(9) The antecedent of the formula in (8) is a true Σ^b_1-sentence of size $2^{O(\delta \cdot n)}$ and has a proof in S^1_2 (and hence in T) of polynomial size, i.e. of size $2^{O(\delta \cdot n)}$.

(10) Applying modus ponens to the formulas in (8) and (9) we get a proof of size $2^{O(\delta \cdot n)}$ the sentence:

$$BigTaut(\lceil C_N \rceil) \ .$$

(11) We claim that the implication:

$$BigTaut(\lceil C_N \rceil) \ \rightarrow \ A(\underline{N})$$

has an S^1_2-proof of size $n^{O(1)}$. This is because it can be obtained from an instance of a universal implication (provable in S^1_2 by Lemma 1.1)

$$BigTaut(g(x)) \ \rightarrow \ A(x)$$

by substituting for x the numeral $\underline{N}$, and by proving $\lceil C_N \rceil = g(\underline{N})$: That is a true Σ^b_1-sentence of size $n^{O(1)}$ and has a polynomial size proof in S^1_2.

(12) Putting (10) and (11) together we get a size $2^{O(\delta \cdot n)}$ proof in T of $A(\underline{N})$. Taking $\delta > 0$ small enough so that $2^{O(\delta \cdot n)} < N$ (this we can do as the O-constant is independent of δ) we get a contradiction with Theorem 0.1.

q.e.d.

Let us remark that instead of using $\mathcal{E}$ and circuit size $2^{\Omega(n)}$ in (i) we could have used $Time(t(n))$ and circuit size $t(n)^{\Omega(1)}$, as long as there is no polynomial upper bound for $t(n)$. This follows by a padding argument or by a simple change to the proof above: Use the diagonal formula for term $s(x) := t(|x|)$ instead of $s(x) := x$ in the step (3).

3 A variant of the theorem

It is not difficult to see that the property of a string to be the truth table of a function on $\{0,1\}^n$ with $2^{\delta \cdot n}$ circuit complexity, or even $2^{\delta \cdot n}$ hard on average, is in the polynomial time hierarchy $\mathcal{PH}$. Taking the lexicographically first such strings (at least one exists of each length 2^n, $n >> 0$, by a simple counting) we see that there is such an f computable in $\mathcal{E}^{\mathcal{PH}}$.

If $\mathcal{NP} = co\mathcal{NP}$ then such an f is in $\mathcal{E}^{\mathcal{NP} \cap co\mathcal{NP}} = \mathcal{NE} \cap co\mathcal{NE}$. If, in addition, $\mathcal{E} = \mathcal{NE}$ then such an f is in $\mathcal{E}$.

This simple (apparently folklore) argument yields the following theorem.[1] We shall use only Part 2 in Section 4 but we state also Part 1 as it is a weaker version (if hardness on average is replaced by circuit size) of Theorem 2.1 actually: It is known, by [11], that the non-existence of a p-optimal proof system implies that $Time(t(n)) \neq NTime(t(n))$ (as long as $t(n) \leq 2^{n^{O(1)}}$) and also $\mathcal{E} \neq \mathcal{NE}$ (the opposite implication is unknown but Verbitsky [18] constructed a relativized world where it does not hold).

Theorem 3.1 *The following two propositions hold.*

1. *At least one of the following three statements is true:*
   - (a) *There is a function $f : \{0,1\}^* \to \{0,1\}$ computable in $\mathcal{E}$ that has $2^{\Omega(n)}$ hardness on average.*
   - (b) $\mathcal{NP} \neq co\mathcal{NP}$.
   - (c) $\mathcal{E} \neq \mathcal{NE}$.
2. *At least one of the following two statements is true:*
   - (a) *There is a function $f : \{0,1\}^* \to \{0,1\}$ computable in $\mathcal{NE} \cap co\mathcal{NE}$ that has $2^{\Omega(n)}$ hardness on average.*
   - (b) $\mathcal{NP} \neq co\mathcal{NP}$.

[1] This argument has been pointed out to me by E. Jeřábek, and has been also noted by V. Kabanets, and replaces my original proof: A simple modification of the proof of Theorem 2.1 shows that $\mathcal{NP} = co\mathcal{NP}$ implies that $\mathcal{NE} \cap co\mathcal{NE}$ contains a function with exponential circuit complexity which was then turned into a function with exponential hardness on average by the construction from [4].

4 Proof complexity generators

By a proof complexity generator we mean a map $g : \{0,1\}^n \to \{0,1\}^m$, $m = m(n)$ and $m > n$, whose bits can be computed in $NTime(m(n)^{O(1)}) \cap coNTime(m(n)^{O(1)})$. This assumption about the computability of the bits is the weakest one allowing to write down, for any $b \in \{0,1\}^m$, a size $m^{O(1)}$ propositional formula $\tau_b(g)$ that is a tautology iff $b \notin Rng(g)$, see below. (The notation $\tau_b(g)$ is somewhat misleading as the formula depends on a particular definition of g and not only on g, but we will ignore this here: The lower bounds conjectured later should hold for all $NTime(m(n)^{O(1)}) \cap coNTime(m(n)^{O(1)})$-definitions.)

A generator is good if for a strong proof system P and with high probability in choosing a random $b \in \{0,1\}^m$, the formula $\tau_b(g)$ requires very long (in particular, super-polynomial) P-proofs. The quality of the generator is measured by the strength of P and by the probability that b yields a hard τ-formula. At present it is not ruled out that some generator g works for all P and all b. Following [17] we shall say that generator g *is hard for* P if all $\tau_b(g)$, for all b's, require super-polynomial size P-proofs (cf.[17]).

The τ-formulas have been defined in [6] and independently in [1], and their theory has made first steps in [7, 16, 8, 17]. I shall not describe the development of the ideas and known lower bound results; this can be found in the introductions to [8] or [17]. Instead I shall briefly describe one motivation and why we speak about "generators".

Proving lower bounds for strong propositional proof system appears hard. In fact, we do not know any such lower bounds. A factor contributing to this is that it is actually not easy to come up with sensible tautologies that would be good candidates for requiring long proofs even in strong systems (cf. [8] for a detailed discussion). The τ-formulas seem to be candidates worth studying in this context.

The word generator is used because some of the usual pseudo-random number generators seem to be good candidates. In particular, a good proof complexity generator must behave as a hitting set generator w.r.t. $\mathcal{NP}/poly$-test (cf.[8]).

Similarly as the existence of good pseudo-random generators can be proved under some computational hardness assumptions (cf.[12, 4]) we may try to reduce the existence of good proof complexity generators to a suitable computational hardness assumption too. This is discussed in [9] with a broader perspective and in the introduction to [17].

The most studied map in this context is the classic Nisan-Wigderson generator, cf. [12]. This has been proposed as possibly a good proof complexity generator in [1] and taken up in [8], although the motivations (and, more importantly, the choice of parameters in the construction and the formalization of the notion of *hardness* of the generators[2]) are different. Let us first recall the definition of the NW-generators (and fix the notation in the process).

Let A be an $m \times n$ 0-1 matrix with ℓ ones per row. $J_i(A) := \{j \leq n \mid A_{ij} = 1\}$. Let $f : \{0,1\}^\ell \to \{0,1\}$ be a boolean function. $NW_{A,f} : \{0,1\}^n \to \{0,1\}^m$ is the NW-generator based on A and f: The i-th bit of output is computed by f from the bits of the input that belong to $J_i(A)$.

Assume that $f \in NTime(t(n)) \cap coNTime(t(n))$. Given particular $NTime(t(n))$-definitions α: $\exists v, \alpha_\epsilon(u,v)$ ($|u| = \ell$, $|v| \leq t(n)$ and α_ϵ p-time) of $f(u) = \epsilon$, for $\epsilon = 0,1$, the τ-formulas are defined by:

$$\tau_b^\alpha \ := \ \bigvee_{i \leq m} \neg\alpha_{b_i}(x \downarrow J_i(A), v^i)$$

where $b = (b_1, \dots, b_m) \in \{0,1\}^m$, x is an n-tuple of variables and v^i are disjoint $t(n)$-tuples of variables. Clearly $\tau_b \in TAUT$ iff $b \notin Rng(NW_{A,f})$. Note that the size of τ_b^α is $t(n)^{O(1)} \cdot m(n)$, and hence we get a size $m^{O(1)}$ formula as long as $t(n) \leq m(n)^{O(1)}$.

An idea, formulated in [1] in general terms and then quite specifically in [17], is that $NW_{A,f}$ forms a good proof complexity generator, as long as A has suitable combinatorial properties (being an (ℓ, d) combinatorial design in the sense of [12]: $J_i(A)$'s have size ℓ and the intersection of any two different rows has size $\leq d$) and as long as f is computationally hard. Specifically, Razborov [17] has made the following conjecture.

Conjecture 4.1 (A. A. Razborov[17, Conjecture 2])
Any NW-generator based on a matrix A which is a combinatorial design with the same parameters as in [12] and on any function f in $\mathcal{NP} \cap co\mathcal{NP}$ *that is hard on average for* $\mathcal{P}/poly$*, is hard for EF.*

Let us interpret the specifications. The parameters in the main construction of combinatorial designs in [12, L.2.5] satisfy: $d = \log(m)$, $\log(m) \leq$

[2] I shall describe neither the set-up from [8] nor the conjectured hardness in terms of pseudo-surjectivity here as it is not relevant to the topic of this paper.

$\ell \leq m$ and $n = O(\ell^2)$.[3] Writing this dually:

$$(1) \qquad m = 2^{\epsilon \cdot n^{1/2}} \ \text{ and } \ \ell = \epsilon \cdot n^{1/2}$$

where $\epsilon > 0$ is a constant (determined by the O-constant in the expression $n = O(\ell^2)$). We are taking the maximal value allowed for m by [12, L.2.5] as that is the chief case in [12], and it also links with [16, 17] studying "function" generators.

The phrase "hard on average for $\mathcal{P}/poly$" presumably means that the hardness of f in the sense of [12] is exponential, similarly as it is needed in [12]:

$$(2) \qquad H_f(\ell) \geq 2^{\Omega(\ell)} \ .$$

The conjecture requires that $f \in \mathcal{NP} \cap co\mathcal{NP}$. We shall relax this condition to:

$$(3) \qquad f \ \in \ NTime(2^{O(\ell)}) \cap coNTime(2^{O(\ell)}) \ .$$

By (1) we have $2^{O(\ell)} = m^{O(1)}$ so (3) means that f is in $NTime(m^{O(1)}) \cap coNTime(m^{O(1)})$. By the discussion earlier this means that the size of the τ-formula will be still $m^{O(1)}$. Thus this modification of the original formulation of the conjecture seems quite harmless and, moreover, very much in the spirit of [12] allowing to compute the output bits of the generator in time $m^{O(1)}$ rather than just $n^{O(1)}$.

Nevertheless, (3) is a modification of the original specifications and so we do not want to talk about Conjecture 4.1 in the next theorem. For this reason let us formulate the following **Statement R**:

R Let g be an NW-generator based on an $m \times n$ matrix A that is an $(\ell, \log(m))$ combinatorial design and on any function f such that the constrains in (1), (2) and (3) are satisfied. Then g is hard for EF.

Theorem 4.2 *Assume that Statement R is true. Then EF is not p-bounded.*

Proof :

Assume that EF is p-bounded. We shall arrive at a contradiction with Statement R. If EF is p-bounded then, in particular, $\mathcal{NP} = co\mathcal{NP}$.

[3] There is also a construction specific for the application to $\mathcal{BPP}$ with parameters satisfying: $\ell = c \cdot \log(m)$ and $n = O(c^2 \cdot \log(m))$, where c is a constant, cf. [12, L.2.6]. That is not good in proof complexity as a trivial proof of the τ-formulas going through all possible seeds would have a polynomial size.

By Theorem 3.1 (Part 2) there is a function f in $\mathcal{NE} \cap co\mathcal{NE}$ that has exponential hardness on average. Having such f we may apply Statement R (suitable matrices A are constructed in [12]) to conclude that EF is not p-bounded. That is a contradiction.

q.e.d.

Note that if we postulate smaller m then although the size of the formula might not be polynomial in m anymore it will still be superpolynomially smaller than $2^{O(n)}$, i.e. than a lower bound to the size of the trivial proof going through all seeds.

An optimist may conclude that it only takes to prove a conditional statement in order to prove that EF is not p-bounded. A pessimist may conclude that if the conclusion of R holds even without its hypothesis then the hypothesis is irrelevant. I think an interesting modification of R may be obtained by stating it for one particular function f, e.g. the hard bit of the discrete logarithm, and maybe claiming the lower bound for $\tau_b(g)$ for a random b with high probability, rather than for all b's.

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