Approximate counting and NP search problems

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November 29, 2021

Abstract

We study a new class of NP search problems, those which can be proved total using standard combinatorial reasoning based on approximate counting. Our model for this kind of reasoning is the bounded arithmetic theory APC₂ of [Jeřábek 2009]. In particular, the Ramsey and weak pigeonhole search problems lie in the new class. We give a purely computational characterization of this class and show that, relative to an oracle, it does not contain the problem CPLS, a strengthening of PLS.

As CPLS is provably total in the theory T₂², this shows that APC₂ does not prove every ∀Σ₁¹ sentence which is provable in bounded arithmetic. This answers the question posed in [Buss, Kołodziejczyk, Thapen 2014] and represents some progress in the programme of separating the levels of the bounded arithmetic hierarchy by low-complexity sentences.

Our main technical tool is an extension of the “fixing lemma” from [Pudlák, Thapen 2017], a form of switching lemma, which we use to show that a random partial oracle from a certain distribution will, with high probability, determine an entire computation of a P^NP oracle machine.

The introduction to the paper is intended to make the statements and context of the results accessible to someone unfamiliar with NP search problems or with bounded arithmetic.

1 Introduction

An NP search problem is specified by a polynomial-time relation R(x, y) and a polynomial p(x). Given an input x, a solution to the problem is any y such that R(x, y) holds and |y| < p(|x|) (where |x| is the length of a string x). We only consider total problems, where a solution is guaranteed to exist for all x. The class of all such problems is called TFNP, standing for total functional NP [31].

Subclasses of TFNP are sometimes described as consisting of all search problems which can be proved to be total by some particular combinatorial lemma or style of argument [31] [33]. For example, the class PPA “is based on the lemma that every graph has an even number of odd-degree nodes” [3]. Often, the particular lemma or argument can be represented by a specific axiomatic theory. In

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this paper we study the class, which we call APPROX, of problems that can be
proved total using reasoning based on approximate counting. Our model of this
kind of reasoning is the theory APC\(_2\) developed by Jeřábek in \([20, 21]\),
which provides machinery to count the size of a set well enough to distinguish
between sets of size \(a\) and \((1 + \varepsilon)a\), for \(a\) given in binary (but not
between sets of size \(a\) and \(a + 1\)), and to formalize a certain amount of induction in this language. In
this way it can carry out the standard proofs of, for example, the finite Ramsey
theorem and the tournament principle \([21]\). We give a non-logical characteriza-
tion of APPROX, as the class of TFNP problems reducible in a certain se-

APC\(_2\) lies between \(T^1\) and \(T^3\). In \([11]\) we pointed out that the NP search
problems typically used in arguments separating \(T^2\) from \(T^2\) for \(i \geq 2\)
could be proved total using approximate counting. This led us to state the fol-

Our result in this paper implies that the answer is yes. The totality of CPLS,
which is provable in \(T^2\), is not provable in APC\(_2\), and thus \(T^3\) proves strictly
more NP search problems total than APC\(_2\) does. This makes APC\(_2\) one of the
strongest natural theories that has been separated from theories higher up in
the bounded arithmetic hierarchy – in fact, from \(T^2\) for the lowest possible \(i\) –
in terms of NP search problems. Intuitively speaking, the conclusion is that the
power of \(T^2\), \(T^3\), . . . to prove many NP search problems total is based on more
than just a limited ability to count.

While our motivation is from logic, an important part of the methods we use
are complexity-theoretic, and may be independently interesting to complexity
theorists. Our main technical tool is the “fixing lemma” from \([42]\). This is
related to the switching lemma of H˚astad \([19]\), which is used in complexity
to separate depth \(d + 1\) circuits from depth \(d\) circuits. The fixing lemma is a
simplified, but more widely applicable, version of this result. It shows that a
random partial assignment can, with high probability, determine the value of
a CNF. We strengthen it slightly, to show that a random partial oracle can
determine an entire computation of a \( P^{NP} \) oracle machine. The proof of our version is almost identical, and many definitions are identical, to what appears in \cite{42}. In the more technical parts of Sections 4-6 will assume the reader has access to that paper.

In the rest of this introductory section we give an overview of bounded arithmetic and the theory \( APC_2 \), describe the structure of \( TFNP \) from this point of view, and outline how we handle relativization and reductions. The paper is then structured as follows.

Section 2. We define CPLS and our search-problem class APPROX, formally state our main Theorems \([12]\) and \([13]\) obtain some corollaries, and give an outline of the proofs.

Section 3. We prove Theorem \([12]\) that the class APPROX captures the \( \forall \Sigma^b_1 \) consequences of \( APC_2 \). This section contains the technical work in logic.

Section 4. We state and prove our version of the fixing lemma.

Section 5. We use the fixing lemma to show that a \( P^{NP} \) computation is determined by a random partial oracle, and derive Theorem \([13]\) that CPLS is not in APPROX.

Section 6. We briefly sketch an alternative way to prove our main result about bounded arithmetic, that \( APC_2 \) does not prove the totality of CPLS, by going through propositional proof complexity rather than NP search problems.

Section 7. We mention some open problems.

Acknowledgements. The first author was partially supported first by grant 2013/09/B/ST1/04390 and then by grant 2017/27/B/ST1/01951 of the National Science Centre, Poland. The second author was partially supported by GAČR project 19-05497S. Part of this research was carried out during the first author’s visit to Prague in 2017, funded by ERC grant 339691. The Institute of Mathematics of the Czech Academy of Sciences is supported by RVO:67985840.

We are grateful to Pavel Pudlák and Pavel Hubáček for discussions about this work.

1.1 Bounded arithmetic

Fix a language \( L_{PV} \) containing a symbol for every function or relation computed by a polynomial-time machine. Then a total NP search problem can be identified with a true \( L_{PV} \) sentence of the form \( \forall x \exists y < t(x) R(x, y) \), where \( R \) is a polynomial-time relation, \( t \) is a polynomial-time function, and \( x \) and \( y \) range over natural numbers written in binary notation. Let \( T \) be any sound theory. The set of such sentences provable in \( T \) then defines a class of search problems. For the class to have some reasonable properties, \( T \) should not be too weak; and to get classes of the kind usually studied in complexity theory, it should not be too strong.

Natural examples of suitable theories \( T \) come from bounded arithmetic, which has close ties to computational complexity. For the purposes of this paper, we will take such theories to be given by a base theory fixing some basic
properties of the symbols of $L_{PV}$, together with one or more axiom schemes that allow us to do stronger kinds of reasoning, typically induction. All axioms are universal closures of bounded formulas, that is, formulas in which all quantifiers appear in the form $\forall x < t$ or $\exists x < t$.

In more detail, a PV formula is a quantifier-free formula of $L_{PV}$. A $\Sigma^b_i$ formula is one of the form

$$\exists x_1 < t_1(\bar{z}) \forall x_2 < t_2(\bar{z}, x_1) \ldots \varphi(\bar{z}, \bar{x})$$

where $\varphi$ is a PV formula, the bounds $t_j$ are $L_{PV}$-terms, quantifiers may appear in alternating $\exists$ and $\forall$ blocks, and there are at most $i$ blocks. Such formulas define precisely the $\Sigma^p_i$ relations, that is, those at level $i$ in the polynomial hierarchy. The $\Pi^b_i$ formulas are defined dually. The universal closure of a formula $\varphi(\bar{z})$ is the sentence $\forall \bar{z} \varphi(\bar{z})$. Given a class of formulas $\Gamma$, we write $\forall \Gamma$ for the set of universal closures of formulas from $\Gamma$. Thus, for instance, a $\forall$ PV sentence states that some polynomial-time computable property holds for all inputs.

We will consider two base theories, both containing only $\forall$ PV sentences. The first and more usual one is the theory $PV$, which comes from Cobham’s characterization of the polynomial-time functions as a function algebra [16, 15]. We will not define $PV$ here, as the details are not important, but it can be thought of as a minimal theory in which all polynomial-time functions are well-behaved.

The second, stronger but simpler theory, which we denote $\forall PV(\mathbb{N})$, consists simply of all $\forall$ PV sentences which are true under the standard interpretation in $\mathbb{N}$. This is simpler to understand than $PV$ because there is no list of axioms to keep in mind, and also works more naturally for defining NP search problems. Our results translate easily between the two, and a reader unfamiliar with bounded arithmetic will not go very wrong by reading $PV$ as $\forall PV(\mathbb{N})$ throughout – see Subsection 1.3.

The important family of theories $T^2_i$, for $i \geq 0$, is defined as

$$T^2_i := PV + \Sigma^b_i \text{-IND}$$

where $\Sigma^b_i \text{-IND}$ is the usual induction scheme for $\Sigma^b_i$ formulas with parameters. $PV$ already proves induction for quantifier-free formulas (as that sort of induction can be witnessed by polynomial-time binary search) so in this setting $T^0_2$ is the same as $PV$. We write $T_2$ for the union of this family.

We can now state a fundamental theorem. By the $\Sigma^b_i$-definable functions of a theory we mean the functions with $\Sigma^b_i$ graphs which the theory proves are total.

**Theorem 1** ([9]). For $i \geq 0$, the $\Sigma^b_{i+1}$-definable functions of $T^2_i$ are precisely the $P^{\Sigma^p_i}$ functions, that is, those that are polynomial-time computable with an oracle from level $i$ of the polynomial hierarchy.

We will also use a related family of theories $S^2_i$, for $i \geq 1$, which is defined by replacing the $\Sigma^b_i$-IND scheme in $T^2_2$ with the apparently weaker scheme $\Sigma^b_i$-LIND.
in which inductions can only run for polynomially many steps (in the binary length of a parameter). We have \( T_2^i \subseteq S_2^{i+1} \subseteq T_2^{i+1} \). Theorem 1 remains true if \( T_2^i \) is replaced by \( S_2^{i+1} \) and/or the base theory PV is replaced by \( \forall PV(N) \).

In addition to being related to computational complexity by Theorem 1, bounded arithmetic is a natural environment in which to ask questions about the provability or consistency of theorems or conjectures from complexity theory. For recent examples see [38, 34].

We now make the above definitions slightly more complicated. As in complexity theory, we typically cannot expect to show that two theories of bounded arithmetic are distinct without either making some extra assumption or working relative to some oracle. We will use oracles. We redefine \( L_{PV} \) to include a unary relation symbol \( \alpha \) standing for “an arbitrary oracle”, and function and relation symbols for all polynomial-time machines with oracle access to \( \alpha \). Other formula classes and theories are redefined to use this extended language. In particular, \( \forall PV(N) \) becomes the set of \( \forall PV \) sentences which are true in \( \langle N, A \rangle \) for every oracle \( A \) interpreting the symbol \( \alpha \). Strictly speaking, we should change the names to \( PV(\alpha), \Sigma^b_k(\alpha), T_i^2(\alpha) \) etc. However, since we never use the relativized versions, we simplify notation by keeping the old names. The results mentioned above still hold.

**An open problem.** By adapting oracle separation results for the polynomial hierarchy, it has been shown that the strength of the (relativized) theories \( T_i^2 \) increases strictly with \( i \): for each \( i \) there is a \( \forall \Sigma^b_{i+1} \) sentence provable in \( T_2^{i+1} \) but not in \( T_2^i \). A pressing open question in proof complexity is whether this remains true if we measure the strength of theories only by their \( \forall \Sigma^b_k \) consequences for some fixed \( k \), in particular for \( k = 1 \).

We write \( \forall \Sigma^b_k(T) \) for the \( \forall \Sigma^b_k \) consequences of a theory \( T \). In particular \( \forall \Sigma^b_1(T) \) is a class of total NP search problems (as long as \( T \) is sound). From [25, 14] we know that

\[
\forall \Sigma^b_1(PV) \subseteq \forall \Sigma^b_1(T_2^1) \subseteq \forall \Sigma^b_1(T_2^2)
\]

and from [14, 43] we know that for any \( i, k \geq 1 \),

\[
\text{if } \forall \Sigma^b_k(T_2^i) = \forall \Sigma^b_k(T_2^{i+1}) \text{ then } \forall \Sigma^b_k(T_2^i) = \forall \Sigma^b_k(T_2^2).
\]

The following is open for \( k \leq 2 \):

\[
\text{does } \forall \Sigma^b_k(T_2^2) = \forall \Sigma^b_k(T_2) ?
\]

The answer is expected to be negative, even for \( k = 0 \), by analogy with the \( \Pi_1 \) separation between \( \Sigma_1 \) and \( \Pi_{i+1} \) given by the second incompleteness theorem. The case \( k = 1 \) seems to be particularly approachable, as classes \( \forall \Sigma^b_1(T) \) have a natural computational interpretation in terms of NP search problems.

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1This is essentially equivalent to a question in propositional proof complexity about separating bounded-depth Frege systems by formulas of fixed depth, and in particular finding a family of small-width CNF’s which have short refutations in bounded-depth Frege but require long refutations in Res(log).
Approximate counting. Jeřábek [20, 21] developed a bounded arithmetic theory for approximate counting. Following [11] we call this theory APC$_2$ and define it as $T^1_2$ together with the surjective weak pigeonhole principle (sWPHP) for $\text{P}^{\text{NP}}$ functions, which asserts that no such function can be a surjection from $n$ to $2n$, for any $n > 0$. APC$_2$ can formalize many arguments in finite combinatorics that use approximate counting, such as the standard proofs of the finite Ramsey theorem and the tournament principle, as well as some probabilistic reasoning. It lies between $T^1_2$ and $T^3_2$ in strength, as this instance of the weak pigeonhole principle is provable in $T^3_2$.

In [11] we asked the analogue of question (1) for APC$_2$ in place of $T^2_2$. More specifically:

$$\forall \Sigma^b_1(\text{APC}_2) = \forall \Sigma^b_1(\text{T}_2) ?$$  (2)

We expected the answer to be “no”, but the opposite did not seem completely implausible. Approximate counting is a powerful tool in finite combinatorics, and typical combinatorially natural examples of hard $\forall \Sigma^b_1$ statements that have been used to separate $T^1_2$ from $T_2$ were known to be provable in APC$_2$ [11]. Moreover it was shown in [12], by formalizing Toda’s theorem, that all of bounded arithmetic collapses to the analogue of APC$_2$ if we add a parity quantifier to the language.

Both [11] and later [2] showed unprovability results for various natural sub-theories of APC$_2$, but these fell well short of answering (2). In fact, they were obtained using a $\forall \Sigma^b_1$ sentence that is actually provable in APC$_2$.

1.2 TFNP

As already discussed, in our language a total NP search problem is simply a true $\forall \Sigma^b_1$ sentence, that is, one of the form $\forall x \exists y < t(x) R(x, y)$ where $R(x, y)$ is a PV formula and $t$ is an $L_{PV}$-term. This represents the search-task of finding a witness $y$, given $x$. We will often assume that the bound $y < t(x)$ is implicit in $R(x, y)$, and we will usually write $R(x, y)$ or just $R$ as a name for the search problem.

As before, polynomial-time is defined relative to an oracle symbol $\alpha$, and we will occasionally use notation like $R(x, y; \alpha)$ to emphasize the specific oracle being used. The oracle leads to a slight complication in what we mean when we say a search problem is total: $\forall x \exists y < t(x) R(x, y; \alpha)$ must be true in $\langle \mathbb{N}, A \rangle$ for every oracle $A$ interpreting $\alpha$. This behaviour is essentially the same as what is called a total type-2 NP search problem in e.g. [3, 10].

We define two important notions of reducibility between search problems $Q$ and $R$. We will introduce one more in Subsection 2.2.

Definition 2. $Q(x, y)$ is polynomial-time many-one reducible, or simply reducible, to $R(x', y')$, written $Q \leq R$, if there are polynomial-time functions $f$ and $g$...
and \(g\) and a polynomial-time relation \(P\) (all of which may query the oracle \(\alpha\)) such that

\[
R(f(x), y' ; P(x, \cdot)) \to Q(x, g(x, y') ; \alpha)
\]

holds for all \(x, y'\) and \(\alpha\), where \(P(x, \cdot)\) represents the oracle \(\{z : P(x, z)\}\). Two problems are equivalent if they are reducible to one another.

**Definition 3.** \(Q(x, y)\) is polynomial-time Turing-reducible to \(R(x', y')\) if there is a polynomial-time relation \(P\) and a polynomial-time oracle machine \(M\) which, on input \(x\), makes a series of (adaptive) queries to \(R(x', y' ; P(\langle x, x' \rangle, \cdot)))\). If all replies are correct, then \(M\) outputs some \(y\) such that \(Q(x, y; \alpha)\).

We are interested in search problem classes corresponding to bounded arithmetic theories, in the sense that the class captures the search problems proved total by the theory (but see Section 1.3 below). There has been a research programme, motivated partly by the logical separation question discussed above, to characterize these classes.

- **PV** corresponds to FP, the class of search problems which can be solved in deterministic polynomial time [16, 9].
- **T\(_1^1\)** corresponds to PLS [23, 13]. A PLS problem is given by polynomial-time neighbourhood and cost functions \(N_x\) and \(C_x\) and domain predicate \(F_x\), such that \(0 \in F_x\) and if \(y \in F_x\), then \(|y| \leq |x|^k\) for some fixed \(k\). A solution to an instance \(x\) is any \(y \in F_x\) such that either \(N_x(y) \notin F_x\) or \(C_x(N_x(y)) \geq C_x(y)\). Such a \(y\) exists because costs cannot decrease indefinitely. A complete problem for the class is to find a local minimum for a function on a bounded-degree graph.
- **T\(_2^1\)** corresponds to CPLS [28], a generalization of PLS described below.
- For \(k \geq 1\), **T\(_k^1\)** corresponds to a class \(G\_k\) defined by the \(k\)-turn game induction principle [43] (see also [5, 6]). Roughly speaking, a complete problem for \(G\_k\) is: given a sequence of \(k\)-round 2-player games, winning strategies for opposite players in the first and last games, and reductions between neighbouring games in the sequence, find an error in one of the strategies or one of the reductions. Equivalent search problems include further generalizations of PLS and principles about feasible Nash equilibria [41], and LLI\(_k\), the \(k\)-round linear local improvement principle [24].

The theory \(T\(_2^1\)\) is equivalent to a natural formalization of “every P\(^{\Sigma^p_2}\) machine has a computation on every input”, essentially by Theorem 1. The search problems above can thus be thought of as the projections onto TFNP of increasingly strong computation models. This can be taken further: there are two “second-order” bounded arithmetic theories, \(U\(_2^1\)\) and \(V\(_2^1\)\), which are equivalent (with respect to their \(\forall \Sigma^b_1\) consequences) to similar statements about computations of, respectively, PSPACE and EXPTIME machines [9, 24]. In terms of NP search problems, from [24, 1] we have:
Figure 1: A diagram showing some inclusions between selected classes in TFNP, including our results. Some classes are named by their complete problems. Solid arrows are strict inclusions, relative to some oracle. Dotted arrows are inclusions not known to be strict. For separations among PPA, PPP, PPAD, PLS see [3, 32, 8, 10]. For other references see Sections 1.2 and 2.2. CPLS separates GI$_3$ from APPROX and also separates GI$_2$ from HOP and PWPP. WEAKPIGEON separates APPROX from PLS. HOP separates APPROX from PWPP. Solid arrows from PPP to HOP [11] and from PPP to PWPP have been omitted for readability.

- $U^1_2$ corresponds to LLI$_{\log}$, the linear local improvement principle with polynomially many rounds.

- $V^1_2$ corresponds to LI, the local improvement principle.

One can think of the classes $FP \subseteq GI_1 \subseteq GI_2 \subseteq \cdots \subseteq LLI_{\log} \subseteq LI$ as forming a backbone for TFNP, arising natural from the hierarchy of bounded arithmetic theories or of computation models – see Figure 1. This could be extended even beyond bounded arithmetic, to say $\forall \Sigma^0_1(\text{PA})$ or $\forall \Sigma^0_1(\text{ZFC})$ [1] [14], or by using potentially stronger systems of reasoning, as in [17].

However, experience suggests that it is difficult to find any natural “combinatorial” NP search problem that is not already provably total in $U^1_2$, and thus reducible to LLI$_{\log}$. In particular, $U^1_2$ is strong enough to formalize the counting arguments needed to prove the totality of complete problems for the well-known

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\footnote{Unfortunately LLI$_{\log}$ is rather complicated to describe or use. The authors believe that it is equivalent to the simpler game induction principle GI$_k$ which is like the principle GI$_k$ of [33] but with polynomially many rounds.}

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\footnote{A related issue in propositional proof complexity is the difficulty of finding candidates for separating the Frege and extended Frege proof systems [1].}
classes PPA and PPP introduced in \[35\]. Thus, all problems in those classes are reducible to LLI_{log}.

On the other hand the bijective pigeonhole principle, called OntoPIGEON in the search problem literature, is a complete problem for the class PPAD \[8\] which is contained in both PPA and PPP. Standard proof-complexity lower bound arguments for the pigeonhole principle \(27, 39\) show that this problem is not provably total in any theory \(T_k^2\) and not reducible to any GI_k.

Finally let us mention the search problem class PWPP \[22\], based on the injective weak pigeonhole principle search problem WEAKPIGEON, which is contained in GI_2 and PPP but not in PLS \[30, 25\]. We will discuss this and two other search problems, RAMSEY and HOP, in Section 2.

It is open whether the hierarchy GI_2 \(\subseteq\) GI_3 \(\subseteq\) . . . is strict. This is essentially the same problem as the separation of \(\forall\Sigma^b_1(T_i+1)\) from \(\forall\Sigma^b_1(T_i)\) discussed above.

1.3 True and provable reductions

In the previous section, we did not explicitly say what it means for a search problem class to correspond to a theory \(T\). The obvious meaning, that the class is precisely \(\forall\Sigma^1_1(T)\), potentially has a problem. Namely, such a class does not have the desirable property of being closed under many-one reductions, unless the reductions work provably in \(T\). There may even be two PV formulas \(R_1\) and \(R_2\) which “semantically” define the same relation on \(\mathbb{N}\), and thus the same search problem by the usual complexity-theoretic definition, but are such that \(T\) proves that one is total but not the other.

There are two natural ways around this issue. One is to define our class as the closure of \(\forall\Sigma^1_1(T)\) under many-one reductions. The other is to stick to theories \(T\) that contain the set \(\forall\PV(\mathbb{N})\) of all true \(\forall\PV\) sentences, and exploit the fact that the statement that a reduction works is such a sentence.

The next lemma shows that, for theories of the kind we consider, these two approaches have the same result. In this paper we prefer the second one.

Lemma 4 (folklore, see also \[18\]). Let \(Q(x, y)\) be an NP search problem. Let \(T\) be a bounded arithmetic theory containing PV and with axioms closed under substituting polynomial-time relations for oracles. Then (1) and (2) below are equivalent. If \(T\) contains \(S^1_2\), then (3) is also equivalent.

1. \(Q\) is provably total in \(T + \forall\PV(\mathbb{N})\).

2. \(Q \leq R\) for some TFNP problem \(R\) provably total in \(T\).

3. \(Q\) is Turing reducible to a TFNP problem \(R\) provably total in \(T\).

Proof. Suppose (1) holds. We have that \(T + \forall z \varphi(z) \vdash \forall x \exists y Q(x, y)\) for some PV formula \(\varphi\) such that \(\forall z \varphi(z) \in \forall\PV(\mathbb{N})\). Hence \(T \vdash \forall x [\exists z \neg \varphi(z) \lor \exists y Q(x, y)]\). Since \(T\) is a bounded theory, by Parikh’s theorem \[36\] we may add some term \(t(x)\) bounding both existential quantifiers. Therefore \(T \vdash \forall x \exists y < t(x) R(x, y)\) where \(R(x, y)\) is the formula \(\neg \varphi(y) \lor Q(x, y)\). Now \(R\) is an NP search problem,
provable total in $T$, and $Q$ is reducible to $R$ in $\mathbb{N}$ using the identity function, since $\varphi(y)$ is true for every $y$ and every oracle. Hence (2) holds.

Now suppose (2) holds. Then (3) is immediate. For (1), from the definition of a reduction, there are PV function symbols $f, g$ and a relation symbol $P$ such that, for every oracle $A$,

$$\langle \mathbb{N}, A \rangle \models \forall x \forall y' [R(f(x), y'; P(x, \cdot)) \rightarrow Q(x, g(x, y'); A)].$$

Now define the search problem $R^*(x', y'; \alpha) := R(x', y'; P(x, \cdot))$. Then $\forall PV(\mathbb{N})$ proves $\forall x' \exists y' R^*(x', y') \rightarrow \forall x \exists y Q(x, y)$, and by the property of closure under substitution for the oracle, $T$ proves that $R^*$ is total. Hence we have (1).

Lastly we show that (3) implies (1) under the stronger assumption. Turing reducibility means that there is a polynomial-time oracle machine $M$ which, on input $x$, makes oracle queries to $R$ and, if the replies are correct, outputs $y$ such that $Q(x, y)$. Formally, for every $A$, $\langle \mathbb{N}, A \rangle \models \forall x \forall w \varphi(x, w)$ for a PV relation $\varphi(x, w)$ expressing that: if $w$ is a computation of $M$ on input $x$, and every oracle query $x'$ in $w$ has a reply $y'$ in $w$ such that $P(x', y')$, then $Q(x, output(w))$. But $T$ proves that, for all $x$, such a $w$ exists, since $\Sigma^b_1$-LIND is enough to construct $w$ query by query. Hence $T + \forall PV(\mathbb{N}) \models \forall x \exists y Q(x, y)$. \qed

2 Main definitions and results

2.1 Coloured polynomial local search

We study a search problem introduced in \[28\]. We will need several results about it from \[22\], so we take the definition verbatim from there.

Let $a, b, c$ be parameters. Consider a levelled directed graph whose nodes consist of all pairs $(i, x)$ from $[0, a) \times [0, b)$. We refer to $(i, x)$ as node $x$ on level $i$. If $i < a - 1$, this node has a single neighbour in the graph, node $f_i(x)$ on level $i + 1$. Every node in the graph is coloured with some set of colours from $[0, c)$. The principle CPLS, coloured polynomial local search, says that the following three statements cannot all be true:

(i) Node 0 on level 0 has no colours.

(ii) For every node $x$ on every level $i < a - 1$, and for every colour $y$, if the neighbour $f_i(x)$ of $x$ on level $i + 1$ has colour $y$, then $x$ also has colour $y$.

(iii) Every node $x$ on the bottom level $a - 1$ has at least one colour, $u(x)$.

When the parameters $a, b, c$ are universally quantified, CPLS is expressed as a $\forall \Sigma^b_1$ sentence about an oracle $\alpha$ encoding the functions $f_i$ and $u$ and a predicate $G$, where $G_i(x, y)$ means “node $x$ on level $i$ has colour $y$”.

To describe it explicitly as a search problem: the inputs are the parameters $a, b, c$ and a solution is a witness that one of items (i)-(iii) above fails. That is, a colour $y$ such that $G_0(0, y)$; or a node $(i, x)$ and a colour $y$ such that $G_{i+1}(f_i(x), y) \land \neg G_i(x, y)$; or a node $(a - 1, x)$ such that $\neg G_{a-1}(x, u(x))$.
To see that the principle is true, or equivalently that the search problem is total, suppose that (i)-(iii) hold simultaneously. Then we can reach a contradiction by arguing inductively on $i$ that for all $i$, some node on level $i$ has no colours. This argument can be formalized as a proof of CPLS in $T^2_2$. Moreover, this has a kind of converse, in that it is shown in [28] that CPLS is complete for the search-problem class $\forall\Sigma^b_1(T^2_2)$ with respect to many-one reductions.

It is worth pointing out that CPLS is a generalization of PLS in the sense that it simplifies to a PLS-complete problem if the parameters are restricted in a certain way, for instance if we fix the number of colours $c$ to 1. In that situation, given $a,b$, we define the domain $F$ of the PLS problem by putting $(i, x) \in F$ if and only if $\neg G_i(x, 0)$ holds, that is, if $x$ does not have the unique possible colour 0 on level $i$. For $(i, x) \in F$, let $N(i, x) := (i + 1, f_i(x))$ and $C(i, x) := a - i$.

Thus, one way of looking at CPLS is that the parameter $a$ is a bound on possible costs, $b$ is a bound on the number of potential neighbours of a given element, and $c$ is a bound on the number of potential reasons why an element could fail to be in the domain. It can be shown that CPLS becomes a PLS problem whenever one of the three parameters is constrained to be is at most polylogarithmic in the maximum of the two others.

### 2.2 Retraction WPHP and $\Sigma^P_2$ search problems

The retraction weak pigeonhole principle [21] asserts that, for $n > 0$, given two functions $f : n \to 2n$ and $g : 2n \to n$ there must be some $v < 2n$ such that $f(g(v)) \neq v$. It is true, because otherwise simultaneously $f$ would be a surjection and $g$ an injection. If $f$ and $g$ are polynomial time, this principle naturally gives rise to a problem in TFNP. We will be in a situation where $f$ and $g$ are $P^{NP}$, and for this we will define a more complex kind of search problem.

**Definition 5.** A $\Sigma^P_2$ search problem is specified by a coNP relation $R(x, y)$ and a polynomial bound $q$ such that $\forall x \exists y < 2^{q(|x|)} R(x, y)$. We will often assume that the bound $q$ is implicit in $R$ and will not write it. The problem represents the search-task of finding such a $y$, given $x$.

As this definition makes sense outside the context of bounded arithmetic, we have written it in standard complexity-theory notation. But we could alternatively define a $\Sigma^P_2$ search problem as a true $\forall\Sigma^P_2$ sentence, in the style of our syntactical definition of TFNP problems.

A basic example of such a search problem is: for a fixed $P^{NP}$ machine $M$, given an input, find a computation of $M$ on this input. Here we assume that a computation includes witnesses for all NP queries that get the answer YES, so that the property of being a (correct) computation is thus coNP. We specify precisely what we mean, as it will be important in what follows.

**Definition 6.** We define a $\Pi^P_2$ formula “$w$ is a computation of $M$ on input $v$”. The formula interprets $w$ as a sequence $\bar{q}, \bar{r}, \bar{y}$ of respectively NP queries, YES/NO replies and witnesses to replies. It expresses that
1. For each $i$, $q_i$ is the $i$-th query asked by $M$ in a computation on input $v$, assuming the previous replies were $r_1, \ldots, r_{i-1}$.

2. For each $i$, if reply $r_i$ is YES then $y_i$ witnesses this.

3. For all sequences $\bar{z}$ of possible counterexamples, for each $i$, if reply $r_i$ is NO then $z_i$ is not a counterexample to this.

The machine only accesses the oracle $\alpha$ via the NP queries. We say that $w$ is a precomputation of $M$ on input $v$ if it satisfies 1. and 2. above.

Note that being a precomputation of $M$ on a given input is a PV formula, so it makes sense to speak of precomputations also when $\alpha$ is only partially defined (as long as the defined part is large enough to verify 2. above). Note also that it is implicit in clause 1. of the definition that each query asked in a computation of $M$ depends only on the input and the previous YES/NO replies to queries, not on the witnesses to the previous replies.

We now give our main definitions.

Definition 7. $r$WP$\Sigma^p_2$ is a class of $\Sigma^p_2$ search problems. A problem in the class is specified by $P^{NP}$ machines computing functions $f_x(u)$ and $g_x(v)$, where we treat one argument $x$ as a parameter. The functions $f_x$ and $g_x$ are constrained to take values less than $2^x$ and $x$ respectively. An input to the problem is a size parameter $x$. A solution is a pair $\langle v, w \rangle$ such that $v < 2^x$, $w$ is a computation of $f_x(g_x(v))$ in the sense of Definition 6, and the output of $w$ is not $v$.

Definition 8. An NP search problem $Q(x, y)$ is PLS counterexample reducible to a $\Sigma^p_2$ search problem $R(x', y')$ if there is a PLS problem $P(x'', y'')$ and polynomial time functions $d$ and $e$ with the following property: for any $x, y', y''$ such that $P(\langle x, y' \rangle, y'')$, either

1. $d(x, y'')$ witnesses that $R(e(x), y')$ is false, or
2. $Q(x, d(x, y''))$.

As in Definition 2, the oracle called by $R$ is allowed to be a polynomial-time variant of the oracle $\alpha$ called by $Q$. Precisely, there is a polynomial-time relation $A$ querying $\alpha$ such that, in the description above, $R$ queries $A(x, \cdot)$ as its oracle rather than $\alpha$.

Definition 9. The search problem class APPROX consists of all NP search problems which are PLS counterexample reducible to an $r$WP$\Sigma^p_2$ problem.

The definition of PLS counterexample reducibility should be understood as follows. We are given $x$ and want to find $y$ such that $Q(x, y)$. We create an input $e(x)$ to $R$ and are given a purported solution $y'$ for which it is claimed that $R(e(x), y')$ – since $R$ is a $\Sigma^p_2$ search problem, this is a coNP claim which we cannot check directly. We then use $\langle x, y' \rangle$ as input for our PLS problem $P$, and find a solution $y''$. Then either 1. or 2. above holds, that is, either the
coNP claim about $R$ was false and $d(x, y''')$ is a counterexample, or $d(x, y''')$ is a solution to our original problem.

As far as we know, this notion of reducibility has not been studied before. We give some examples. Unfortunately, in the examples we know which are relatively simple, one part or another of the definition becomes trivial.

Firstly, every PLS problem $Q(x, y)$ is PLS counterexample reducible to the trivial $\Sigma^p_2$ search problem defined by letting $R(x', y')$ hold for all $y' < 2^{|x'|}$. In the reduction, both $d$ and $e$ are the identity mapping and $P((x, y'), y''')$ holds exactly if $Q(x, y''')$ does.

Now consider four search problems:

1. The $\Sigma^p_2$ problem TOURNAMENT: given $x$ and a binary oracle $\alpha$ representing a tournament on the set of vertices $[0, x)$, find $y$ coding a set of at most $\lceil \log x \rceil$ vertices which dominates the tournament, that is, coding vertices $y_1, \ldots, y_k$ such that for all $z < x$ we have $\alpha(y_i, z)$ for some $i$.

2. The TFNP problem HOP (the *Herbrandized ordering principle*\(^5\)): given $x$ and oracles for a binary relation $\preceq$ and a unary function $h$, find either a witness that $\preceq$ restricted to $[0, x)$ is not a total ordering, or a witness that $h$ is not the $\preceq$-immediate predecessor function on $[0, x)$.

3. The TFNP problem RAMSEY: given $x$ and an oracle $R$ for a graph on $[0, x)$, find $y$ encoding a set $s \subseteq [0, x)$ of cardinality $\lfloor \log x / 2 \rfloor$ such that $[s]^2$ is homogeneous with respect to $R$.

4. The TFNP problem WEAKPIGEON\(^6\): given $x$ and an oracle $g$ for a function from $[0, 2x)$ to $[0, x)$, find $v, v' < 2x$ such that $v \neq v$ and $g(v) = g(v')$.

The class of problems that are reducible to WEAKPIGEON was given the name PWPP in \[^22\].

HOP is reducible to TOURNAMENT in the following sense. We are given a size $x$ and a binary relation $\preceq$ on $[0, x)$ for which we want to solve HOP. Consider $\preceq$ as a tournament on $[0, x)$ and give it as input to TOURNAMENT. Suppose $s = \{y_1, \ldots, y_k\}$ is a purported solution to TOURNAMENT, with $k$ polylogarithmic in $x$. By polynomial-time search we can find either a witness in $s$ that $\preceq$ is not a total ordering (thus solving HOP), or a $\preceq$-least element $y_i$ of $s$. In the second case, compute $y' := h(y_i)$. If $y' \succ y_i$, then $y_i$ is a solution to HOP. Otherwise, comparing $y'$ to each $y_j \in s$ will either give us a witness that $\preceq$ is not a total ordering or reveal that $y'$ witnesses that $s$ is not a correct solution to TOURNAMENT.

The above is a description of a PLS counterexample reduction of HOP to TOURNAMENT. The function $e$ translating input to HOP into input to TOURNAMENT is the identity. We then compute from $s$ either a solution to HOP

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\(^5\)Herbrandized refers to the presence of the predecessor function $h$ which turns the task of finding the $\preceq$-smallest element of $[0, x)$, which is naturally a $\Sigma^p_2$ problem, into a TFNP problem. The generalized iteration principle of \[^14\] is similar in spirit, as is the graph ordering principle in the propositional proof complexity literature.

\(^6\)This is different in a non-essential way from how this problem is defined in \[^22\], where it takes as input a circuit for the function $g$. 

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or a witness that \( s \) is not a solution to TOURNAMENT. However we can do
this computation in polynomial time, while Definition 8 more generally allows
it to be done by a call to a PLS problem \( P \) (assisted by a polynomial time
"decoding" function \( d \)).

Our final examples of PLS counterexample reducibility are related to the
class APPROX. We will show in Theorem 12 below that this class coincides
with the class of NP search problems that are provably total using approximate
counting. This has the following consequence.

**Corollary 10** (of Theorem 12). The NP search problems HOP, RAMSEY
and WEAKPIGEON are in APPROX. Therefore they are PLS counterexample reducible to rWPHP\(_2\) problems.

**Proof.** The problems RAMSEY [40, 21], HOP [11] and WEAKPIGEON are all
provably total in APC\(_2\). (This is also true for a stronger version of HOP in
which \( \leq \) is only required to be a partial ordering, not a total ordering [11].)

We are only aware of a simple direct proof of the reduction in the case of
WEAKPIGEON. Suppose we are given a polynomial-time function \( g \) and a
parameter \( x \) as input. To solve WEAKPIGEON we want to find a collision
in \( g \), which we interpret as a function from \( 2x \) to \( x \). Let a function \( f \) be given
by a P\(^{NP}\) machine inverting \( g \), as follows. On input \( u < x \), \( f \) queries whether
\( \exists v < 2x \ (g(v) = u) \). If the answer is NO, \( f \) gives up and outputs 0. If the answer
is YES, \( f \) outputs the maximal such \( v \), found by binary search. Consider the
problem in rWPHP\(_2\) specified by \( f \), \( g \) and the parameter \( x \). Recall that a
solution is a pair \( \langle v, w \rangle \) such that \( v < 2x \), \( w \) is a computation of \( f(g(v)) \) in the
sense of Definition 6, and the output of \( w \) is not \( v \).

In the PLS counterexample reduction of WEAKPIGEON to this rWPHP\(_2\)
problem, the auxiliary functions \( d \) and \( e \) will be the identity. The reduction
procedure \( P \) will once again be polynomial-time. To describe it, suppose we are
given \( v \) and a precomputation \( w \) of \( f \) on input \( u = g(v) \), with output \( v' \neq v \) (we
can ignore the computation of \( g(v) \), since \( g \) is polynomial time). We want to
either find a collision in \( g \), or a witness that \( w \) is not a computation of \( f \), that
is, that some NO reply recorded in \( w \) is wrong.

First suppose that \( v' = 0 \) and was output by \( f \) because the reply to the first query "\( \exists z < 2x \ (g(z) = u) ? \)" was NO. Then \( v \) is a witness that this reply
is wrong. Otherwise, \( v' \) was found by binary search. In this case, we follow \( w \)
until the first place where the binary search interval excludes \( v \). Then, if the
interval is strictly below \( v \), we can use \( v \) to witness that the most recent NO reply
was wrong. If it is above \( v \), then this has to be the result of a YES answer, with a witness \( v'' > v \) recorded in \( w \). Thus we have found a collision,
since \( g(v'') = g(v) = u \).

Finally let us record here a fact which appears in Figure 1.

**Proposition 11.** Relative to an oracle, HOP is not reducible to WEAKPIGEON
and therefore is not in PWPP.
This is implicit in the proof in \[11\] that $T^1_2 + \text{WPHP}$ does not prove HOP is total. It explicitly follows from \[33\], which generalizes this result and works over the stronger base theory $\forall \text{PV}(\mathbb{N})$.

2.3 Results

Recall that APPROX was defined in the last subsection as the set of NP search problems which are PLS counterexample reducible to an $\text{rWPHP}_2$ problem.

**Theorem 12.** $\forall \Sigma^0_1(\text{APC}_2 + \forall \text{PV}(\mathbb{N})) = \text{APPROX}$.

In other words, APPROX captures the class of TFNP problems which are provably total using approximate counting. This is proved in Lemmas \[18\] and \[20\] in Section \[3\] by applying standard witnessing techniques from bounded arithmetic to the definition of the theory APC$_2$.

**Theorem 13.** CPLS is not in APPROX.

This is proved in Section \[3\] using a lemma about random oracles proved in Section \[4\]. We briefly sketch the proof. We first fix an alleged PLS counterexample reduction of CPLS to a problem from $\text{rWPHP}_2$ specified by a pair of P$^{\text{NP}}$ functions $f$ and $g$, then choose a large size parameter $n$ and use it to set suitable values for the parameters $a, b, c$ of CPLS. We define a notion of a “legal” partial oracle, which in particular is one which does not contain any witness to CPLS. We adapt a lemma on random restrictions from \[42\] to show that with high probability a random partial oracle $\rho$ from a certain distribution will “fix” YES or NO replies to all NP queries made in a P$^{\text{NP}}$ computation, in the sense that these replies will never become wrong in any legal extension of $\rho$ (Lemma \[30\]). It follows that most partial oracles $\rho$ from this distribution will fix computations of $(f \circ g)(v)$ in this sense on most inputs $v$. This is enough for $\rho$ to determine a solution $(v, w)$ to our instance of $\text{rWPHP}_2$ for which it is difficult to find a counterexample (Lemma \[31\]). Finally we again adapt a proof from \[12\] to show, by an Adversary argument in which the Adversary’s strategy uses only legal extensions of $\rho$, that our PLS reduction is not able to find a witness to CPLS from $(v, w)$.

Our main result about bounded arithmetic, answering the question posed in \[11\], is an immediate consequence of Theorems \[12\] and \[13\]:

**Corollary 14.** The principle CPLS is not provable in APC$_2$. Since it is provable in $T^1_2$, it follows that, in the relativized setting, APC$_2$ does not prove all $\forall \Sigma^0_1$ consequences of full bounded arithmetic $T^1_2$.

This naturally also limits the strength of theories that are provable in APC$_2$, such as the following one based on the usual (non-Herbrandized) ordering principle.

**Corollary 15.** Consider the theory consisting of $T^1_2$ together with axioms stating that for every PV formula $R(x, y)$ and every $a$, if $R$ is a partial ordering on $[0, a)$ then $[0, a)$ contains an $R$-minimal element. This theory, in the relativized setting, is strictly weaker than $T^2_2$. 

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Proof. This theory is provable in $T_2^2$ by straightforward induction on $a$. It is also provable in APC$_2$ by an entirely different proof involving a reduction to the tournament principle, as is shown in [11] (by an argument due to Jeřábek). By Corollary 14 the theory does not prove CPLS, hence is weaker than $T_2^2$. \Box

Both CPLS and the ordering principle have short proofs in the resolution propositional proof system, and the argument above could also be used to show that the ordering principle is not “complete” for resolution, in the sense that there are things with short proofs in resolution which do not follow from it. But it is not clear what the most suitable notion of “follow from” is here.

Corollary 16. CPLS is not polynomial-time Turing reducible to RAMSEY or HOP.

Proof. RAMSEY and HOP are in APPROX by Corollary 10. If CPLS were polynomial-time Turing reducible to either of these problems, then Lemma 4 would imply that CPLS is provable in APC$_2 + \forall \text{PV}(N)$, contradicting Theorem 12 and Theorem 13. \Box

3 Witnessing and definability

This section contains our main technical work in logic: a proof of Theorem 12 via two lemmas corresponding to the two containments in the statement of the theorem. The proofs assume some familiarity with bounded arithmetic.

Intuitively, APC$_2$ is a combination of $T_2^1$ and the weak pigeonhole principle, and what we show is that the NP search problems provably total in APC$_2$ arise as a combination of PLS (which is known to correspond to $T_2^1$ [13]) and the weak pigeonhole principle, with an important difference that, while a proof can make many “calls” to WPHP, our reductions only allow one call. We also introduce the following technical condition on reductions, which we will need in our non-reducibility proof in Section 5.

Definition 17. We say that a NP search problem $Q$ is cleanly PLS counterexample reducible to a $\Sigma_2^p$ search problem $R$, if $Q$ is PLS counterexample reducible to $R$ as in Definition 8 with the extra condition that the function $e$, which produces inputs to $R$ from inputs to $Q$, does not make any oracle calls.

It may be helpful to think of such an $e$ as a translation between size parameters, for which the structure of the oracle does not matter.

Lemma 18. Every NP search problem provably total in APC$_2 + \forall \text{PV}(N)$ is cleanly PLS counterexample reducible to an rWPHP$_2$ problem.

Proof. Let $Q(x, y)$ be an NP search problem. Assume that

\[
\text{APC}_2 + \forall z \varphi(z) \vdash \forall x \exists y Q(x, y),
\]

where $\varphi(z)$ is a PV formula such that $\text{N} \models \forall z \varphi(z)$ for all oracles. Thus we have

\[
\text{APC}_2 \vdash \exists y Q(x, y) \lor \exists z \neg \varphi(z).
\]
Writing out the definition of APC, this means

\[ T^3_2 + \forall b, c \exists v < 2b \forall u < b e(c, u) \neq v \vdash \exists y Q(x, y) \land \exists z \neg \varphi(z) \]

where the formula on the left is sWPHP for a universal P

NP

running code c on input u with time bound \(|c|\). The quantifier \(\forall b, \ldots\) should strictly speaking be \(\forall b \neq 0, \ldots\) but we suppress this here and below for the sake of readability. Replacing \(T^3_2\) with the stronger theory \(S^2_2\) and moving sWPHP to the right hand side gives

\[ S^2_2 \vdash [\exists b, c < s'(x) \forall v < 2b \exists u < b e(c, u) = v] \land \exists y Q(x, y) \land \exists z \neg \varphi(z), \]

(3)

where we have also used Parikh’s theorem \([36]\) to bound \(b\) and \(c\) by some term \(s'(x)\). We may assume \(s'(x)\) has the form \(2^{2|x|^k} + x\) for some \(k \in \mathbb{N}\), where “+x” is included to make it possible to recover \(x\) from \(s'(x)\) as required below.

The formula in square brackets asserts that sWPHP fails for \(e\) for parameters \(b, c\) less than \(s'\). By a standard technical fact stated and proved below as Lemma \([19]\) we can amplify this failure to a \(P^{NP}\) function \(F\) which is a surjection from \((s')^5\) to \(2(s')^5\), with no parameters. This lets us remove one block of existential quantifiers from \([3]\) to give, setting the term \(s\) to be \((s')^5\),

\[ S^2_2 \vdash [\forall v < 2s \exists u < s F(u) = v] \land \exists y Q(x, y) \land \exists z \neg \varphi(z). \]

(4)

To match the definition of rWPHP, we define a \(P^{NP}\) function of two arguments \(a, u\) by \(f_a(u) = \min(F(u), 2a - 1)\). Then \((4)\) is equivalent to

\[ S^2_2 \vdash [\forall v < 2s \exists u < s f_s(u) = v] \land \exists y Q(x, y) \land \exists z \neg \varphi(z). \]

(5)

The sentence in \((5)\) is \(\forall \Sigma^b_2\), so by Buss’ witnessing theorem for \(S^2_2\) \([9]\) there is a \(P^{NP}\) machine which, provably in \(T^3_1\), maps the input parameters \(x, v\) to a triple \(\langle u, y, z\rangle\) witnessing one of the three existential quantifiers. Let \(g\) be defined so that \(g_0(x)(v)\) first computes \(x\) from \(s(x)\) and then outputs the first component \(u\) of this witnessing function applied to \(\langle x, v\rangle\), as long as \(u < s\); otherwise, \(g\) outputs 0. We have

\[ T^3_1 \vdash [\forall v < 2s f_s(g_s(v)) = v] \land \exists y Q(x, y) \land \exists z \neg \varphi(z). \]

Now, \(f_s(g_s(v)) = v\) can be written as a \(\Pi^b_2\) formula:

\[ \forall w [w \text{ is a computation of } f_s(g_s(v)) \rightarrow \text{output}(w) = v], \]

where \(w\) is suitably bounded by a term in \(x\), and “\(w\) is a computation of \(f_s(g_s(v))\)” is a \(\Pi^b_2\) formula as in Definition \([9]\) describing the \(P^{NP}\) machine that first computes \(g\) and then computes \(f\) on the output.

So we have

\[ T^3_2 \vdash \forall v < 2s \forall w [w \text{ is not a computation of } f_s(g_s(v)) \lor \text{output}(w) = v] \land \exists y Q(x, y) \land \exists z \neg \varphi(z). \]
The formula in square brackets is now $\Sigma^0_1$, so by the PLS witnessing theorem for $T^0_2$ \( \text{[13]} \) there is a PLS problem $P(x''', y''')$ witnessing this whole sentence. That is, if we solve $P$ on input $x''' = \langle x, v, w \rangle$ and find $y'''$ such that $P(x''', y''')$, then one of the following holds:

1. $w$ is not a precomputation of $f_s(g_s(v))$, or has output $v$,
2. $y'''$ is a tuple containing a witness that some NO reply in $w$ is wrong,
3. $y'''$ is a tuple containing a witness to $\exists y Q(x, y)$.

We know that $y''$ cannot contain a witness to the last disjunct $\exists z \varphi(z)$ as by assumption $\mathbb{N} \models \forall z \varphi(z)$.

Using the notation of Definition\[S\] letting $d$ be the function that outputs the witness of incorrectness in case 2., and the witness to $\exists y Q(x, y)$ in case 3., and setting $e(x) = s(x)$, we see that $Q$ is cleanly PLS counterexample reducible to the rWPHP problem given by $f$ and $g$.

To complete the proof of Lemma\[15\] we need a technical lemma from \[45\] Section 2] about “amplifying” failures of WPHP.

**Lemma 19.** Let $e$ be a $\text{P}^{\text{NP}}$ function. There is a $\text{P}^{\text{NP}}$ function $F$ such that provably in $S^2_2$, if $e(c, \cdot)$ is a surjection $b \rightarrow 2b$ then $F(a, b, c, \cdot)$ is a surjection $b \rightarrow 2^{\lvert a \rvert}b$. Furthermore there is a $\text{P}^{\text{NP}}$ function $F'$ such that provably in $S^2_2$, if $e(c, \cdot)$ is a surjection $b \rightarrow 2b$ for some $0 < b, c < s$, then $F'$ is a surjection $s^5 \rightarrow 2s^5$ (with no parameters).

**Proof.** Let $h(a, b, c, u)$ be the function that, assuming $u < 2^{\lvert a \rvert}b$, interprets $u$ as a pair $\langle u_0, u_1 \rangle$ with $u_0 < b$, $u_1 < 2^{\lvert a \rvert}$, and outputs the number $e(c, u_0)2^{\lvert a \rvert} + u_1$. If $e(c, \cdot)$ is a surjection from $b$ onto $2b$, then $h(a, b, c, \cdot)$ is a surjection from $2^{\lvert a \rvert}b$ onto $2^{\lvert a \rvert + 1}b$. Consider now the function $F(a, b, c, u)$ which, given $u < b$, applies the composition of $e(c, \cdot), h(1, b, c, \cdot), h(2, b, c, \cdot), \ldots, h(2^{\lvert a \rvert - 1}, b, c, \cdot)$, to $u$, in the order shown. If $e(c, \cdot)$ is a surjection from $b$ onto $2b$, then $F(a, b, c, \cdot)$ is a surjection from $b$ onto $2^{\lvert a \rvert}b$. To prove this, we consider any fixed number $v < 2^{\lvert a \rvert}b$ and show by reverse induction on $i < \lvert a \rvert - 1$ that $v$ can be obtained by applying the composition of $h(2^i, b, c, \cdot), h(2^{i + 1}, b, c, \cdot), \ldots, h(2^{\lvert a \rvert - 1}, b, c, \cdot)$ to some argument $u$ below $2b^i$. This is where we need the scheme $\Sigma^0_2$-LIND, which is available in $S^2_2$.

For the last claim, we note that if we code quadruples as $\langle \langle \cdot, \cdot \rangle, \langle \cdot, \cdot \rangle \rangle$ where $\langle \cdot, \cdot \rangle$ is Cantor’s pairing function, then for each $k$ larger than a fixed natural number any quadruple of numbers less than $k$ has code less than $k^5$. We treat the argument $x < s^5$ of $F'$ as a quadruple $\langle a', b', c', u' \rangle$ and set $F'(\langle a', b', c', u' \rangle) := F((a' + 1)^5, b', c', u')$. Since such arguments $x$ will cover all quadruples of numbers below $s$, it follows from the properties of $F$ that $F'$ contains every number less than $2^6s^5b$ in its range, and in particular every number less than $2s^5$. \( \square \)

The next lemma is the converse to Lemma\[15\]
Lemma 20. If $Q$ is an NP search problem PLS counterexample reducible to an rWPHP$_2$ problem, then $Q$ is provably total in $\text{APC}_2 + \forall \text{PV}(\mathbb{N})$.

Proof. Let $Q$ be an NP search problem that is PLS counterexample reducible to the rWPHP$_2$ problem given by the functions $f$ and $g$. Let $P(x'', y'')$, $d, e$ be as in the definition of PLS counterexample reducibility. Consider the PV formula $\xi(x, v, w, y, y'')$ defined by

$$v < 2e(x) \land P(\langle x, \langle v, w \rangle \rangle, y'') \land y = d(x, y'') \land [y \text{ does not witness that } w \text{ is not a computation of } f_e(g_e(v)) \neq v],$$

with $y''$ suitably bounded. We view $\xi$ as defining an NP search problem with input $x$ and output $\langle v, w, y, y'' \rangle$. Notice that $\xi(x, v, w, y, y'')$ implies that $Q(x, y)$, so we have $Q \leq \xi$.

We claim that $\text{APC}_2$ proves that $\xi$ is total. To see this, work in $\text{APC}_2$ and consider some input $x$. By sWPHP(PV$_2$), there is some $v < 2e(x)$ which is outside of the range of $f_e(x)$ on inputs below $e(x)$. By $T^1$, there is some computation of $f_e(x)(g_e(v))$, say $w$, which by the choice of $v$ must produce an output different from $v$. Again by $T^2$, there is a solution to $P$ on input $\langle x, \langle v, w \rangle \rangle$, say $y''$. Clearly, $y = d(x, y'')$ cannot witness that $w$ is not a computation of $f_e(x)(g_e(x))$ with output different from $v$, so $\xi(x, v, w, y, y'')$ holds. This proves the claim.

We have shown that $Q$ is reducible to a problem provably total in $\text{APC}_2$. It follows from Lemma 4 that $Q$ is provably total in $\text{APC}_2 + \forall \text{PV}(\mathbb{N})$.

Lemmas 18 and 20 have the following additional consequence.

Corollary 21. The class APPROX can equivalently be defined as the set of NP search problems which are cleanly PLS counterexample reducible to an rWPHP$_2$ problem.

4 Fixing lemma

This section contains our main technical result in complexity, Lemma 27, which is an extension of the “fixing lemma” from [12]. There, the fixing lemma is a limited switching lemma which says the following: given suitable parameters $a, b, c$ for CPLS, for a well-chosen probability distribution on partial restrictions to an oracle $\alpha$ encoding $(f_i)_{i < a - 1}$, $u, (G_i)_{i < a}$, a random restriction has a relatively high probability of determining the value of a narrow CNF in propositional variables standing for bits of $\alpha$. Importantly, the restriction does not reveal a witness to CPLS; in particular, the (unsatisfiable) CNF asserting that there is no witness to CPLS has to be determined to be true.

In our application in the proof of Theorem 13, we want to fix answers to the NP queries made in a $\text{P}^\text{NP}$ computation. Each query is (the negation of) a CNF, but now there are many of them, and they are made adaptively depending on earlier replies. So we cannot use the lemma from [12] directly. Instead we adapt
the proof to show that given a low-depth decision tree labelled with CNFs, with high probability a random restriction fixes the truth values of all CNFs along some branch. This is the basic content of Lemma 27 below.

Our definitions are essentially the same as in [42], and so is one of the proofs. We will repeat some definitions almost verbatim, but will only give high-level descriptions of some other definitions and of the unchanged proof details.

We think of the bits of the oracle as propositional variables. So, for example, for each node \((i, x)\) there are \(\log b\) variables \((f_i(x))_0, \ldots, (f_i(x))_{\log b-1}\) expressing the value of \(f_i(x)\). A total oracle is defined by a total assignment to all variables. We will be working with partial oracles, which we will also call partial assignments or restrictions.

We copy in full the definition of a random restriction from [42]. First, a path in a partial assignment \(\beta\) is a maximal sequence \((i, x_0), \ldots, (i+k, x_k)\) of nodes such that \(f_{i+j}(x_j) = x_{j+1}\) in \(\beta\) for each \(j \in [0, k]\). Note that a path may consist of a single node \((i, x)\) if \(x\) is neither in the domain of \(f_i\) nor in the range of \(f_{i-1}\). If all functions \(f_i\) are partial injections, then every node is on some unique path.

**Definition 22.** ([42, Definition 5.7]) Fix parameters \(0 < p, q < 1\). Let \(\mathcal{R}_{p,q}\) be the distribution of random restrictions chosen as follows.

1. For each pair \(i < a\) and \(x < b\), with probability \((1-p)\) include \((i, x)\) in a set \(Z\). For each \(i < a\), choose \(f_i\) uniformly at random from the partial injections from the domain \(\{x < b : (i, x) \in Z\}\) into \(b\).

2. Set colours on the path beginning at \((0, 0)\) so that \(G_i(x, y) = 0\) for all \(y\) for all nodes \((i, x)\) on that path.

3. For every other path \(\pi\), with probability \((1-q)\) colour \(\pi\) randomly with one colour. That is, choose uniformly at random a colour \(y\) and, for every node \((i, x)\) on \(\pi\), set \(G_i(x, y) = 1\) and then set \(G_i(x, y') = 0\) for all \(y' \neq y\).

4. Finally consider each node \((a-1, x)\) on the bottom level. It is on some path \(\pi\). If \(\pi\) was coloured at step R3, then set \(u(x) = y\) where \(y\) is the unique colour assigned to \(\pi\) (that is, \(G_{a-1}(x, y) = 1\)). Otherwise leave \(u(x)\) undefined.

We will also use \(\mathcal{R}_{p,q}\) to denote the support of this distribution.

We take the definitions of legal restrictions and good restrictions from [42, Definition 5.4 and Lemma 5.8]. Legal restrictions are those that meet a minimal standard of "niceness" – on every path either no colour variables are set, or they are all set in one of a few particular ways which do not immediately witness CPLS. Among restrictions \(\rho \in \mathcal{R}_{p,q}\), the only ones that are not legal are those that contain a path connecting \((0, 0)\) with some node of the form \((a-1, x)\). Our lower bound in the next section will make use of a game played between a Prover, who is trying to witness CPLS by making oracle queries, and an Adversary who is trying to answer queries in a way that does not witness CPLS. It will turn
out that the Adversary can restrict herself to answers that come from legal restrictions. In effect, we do not have to worry about the evaluation of formulas under restrictions which are not legal.

A *good* restriction is a legal one which is of typical size, measured in various ways – in particular no path is very long, and there is a reasonable fraction of variables unset at every level. A *bad* restriction is one which is not good.

It may be useful to keep in mind what the analogous definitions would be if we were dealing with the more familiar pigeonhole principle PHP instead of CPLS. A legal restriction would be any restriction representing a partial injection. With probability parameter $p$, a random restriction would choose holes independently with probability $1-p$, and then randomly map some pigeons to the chosen holes. A good restriction would be a legal one that leaves at least, say, a fraction $p/2$ of holes unset.

We choose a suitable large $n$ and fix our parameters as $a = b = n$, $c = \lfloor n^{1/7} \rfloor$, $p = n^{-4/7}$ and $q = n^{-2/7}$, where $b$ and $c$ are powers of 2.

**Lemma 23.** ([42, Lemma 5.8]) The probability that a random restriction is bad is exponentially small in $n$.

**Definition 24.** Let $\rho$ be a restriction. We say that a CNF $B$ is:

- *fixed to 0* by $\rho$ if $\rho$ falsifies $B$, that is, if for some conjunct in $B$ each literal in the conjunct is set to 0 by $\rho$,

- *fixed to 1* by $\rho$ if it is not fixed to 0 by any legal extension of $\rho$.

Note that a legal restriction can fix a CNF to at most one truth value. The following proposition is therefore obvious.

**Proposition 25.** If $\rho$ fixes a CNF $B$ then every extension of $\rho$ also fixes $B$ to the same value.

It follows from the proof of the “fixing lemma” [42, Lemma 5.9] that, for $k$ reasonably small compared to $n$, the probability that a given $k$-CNF is fixed by a random restriction is relatively high – in fact, the probability that it is not fixed is $O(kn^{-1/7})$. We need a slightly more general version that also bounds some conditional probabilities.

**Lemma 26** (conditional fixing lemma). Let $A_1, \ldots, A_m$ be a sequence of $k$-CNFs and $e_1, \ldots, e_m$ be a sequence of 0/1 values such that

$$\Pr[\rho \text{ is bad} \mid \rho \text{ fixes each } A_i \text{ to } e_i] < 1/2.$$  

Let $B$ be a $k$-CNF. Then

$$\Pr[\rho \text{ does not fix } B \mid \rho \text{ is good and } \rho \text{ fixes each } A_i \text{ to } e_i] < 12kn^{-1/7}.$$  

**Proof.** If we remove the CNFs $A_i$, this is essentially the “fixing lemma” of [42, Lemma 5.9] and our proof is almost identical. Define

$$F = \{\rho \in R_{p,q} : \rho \text{ fixes each } A_i \text{ to } e_i\}$$  

$$G = \{\rho \in R_{p,q} : \rho \text{ is good}\}$$  

$$S = \{\rho \in F \cap G : \rho \text{ does not fix } B\}$$

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so that our $S$ is the intersection with $F$ of the set $S$ defined in [42]. The assumption gives us that $\Pr[F \cap G] / \Pr[F] \geq 1/2$ and our goal is to show that $\Pr[S] / \Pr[F \cap G]$ is small.

Every $\rho \in S$ does not falsify $B$ but does have some legal extension which falsifies $B$. Exactly as in [42] we define a function $\theta$ on $S$ by $\theta(\rho) = \sigma'$, where $\sigma'$ is a certain minimal legal extension of $\rho$. We have, over $\rho \in S$,

1. $\Pr[\theta(\rho)]/\Pr[\rho] \geq \frac{1}{2} n^{1/7}$
2. $\theta$ is at most $3k$-to-one
3. $\theta(\rho) \in F$ (although it may happen that $\theta(\rho) \notin F \cap G$).

Items 1 and 2 are proved as in [42]. Item 3 is immediate from Proposition 25.

Now partition $S$ as $S_0, \ldots, S_{3k-1}$ where $S_i = \{\rho \in S : \rho$ is the $i$th preimage of $\theta(\rho)\}$. Then

$$\Pr[S_i] = \sum_{\rho \in S_i} \Pr[\rho] = \sum_{\rho \in S_i} \Pr[\theta(\rho)] \frac{\Pr[\rho]}{\Pr[\theta(\rho)]} \leq 2n^{-1/7} \sum_{\rho \in S_i} \Pr[\theta(\rho)] \leq 2n^{-1/7} \Pr[F]$$

where for the last inequality we use that $\sum_{\rho \in S_i} \Pr[\theta(\rho)] \leq \Pr[F]$, since $\theta$ is an injection from $S_i$ to $F$. This step is the main difference from [42], which uses only that $\theta$ is an injection from $S_i$ to $R_{p,q}$, giving the weaker bound $\sum_{\rho \in S_i} \Pr[\theta(\rho)] \leq \Pr[R_{p,q}] = 1$.

It follows that $\Pr[S] = \Pr[S_0] + \cdots + \Pr[S_{3k-1}] \leq 6kn^{-1/7} \Pr[F]$. Hence $\Pr[S] / \Pr[F \cap G] \leq 12kn^{-1/7}$ as required.

**Lemma 27.** Consider a complete binary decision tree in which each internal node is labelled with a $k$-CNF and has outgoing edges for NO and YES answers. A node $z$ and a restriction $\rho$ are compatible if $\rho$ is good and, for every CNF $B$ on the path down from the root to $z$, $\rho$ fixes $B$ to the value specified by the outgoing edge along the path.

Let $\varepsilon = \Pr[\rho$ is bad$]$. A node $z$ is big if $\Pr[\rho$ is compatible with $z] > \varepsilon$. Let $S_d$ be the set of good restrictions $\rho$ which are compatible with some big node at depth $d$. Then

$$\Pr[S_d] \geq 1 - d \cdot 12kn^{-1/7} - 2^{d+1}\varepsilon.$$

This will be used in the next section, where the decision tree will model a computation of a $P^{NP}$ machine. In particular $d$ and $k$ will be polylogarithmic in $n$ and $\varepsilon$ will be exponentially small in $n$. It follows from the lemma that at least one node on the bottom level, and thus at least one computation of the machine, is compatible with some $\rho$.

**Proof.** We use induction on $d$. For the base case $d = 0$, first observe that every good restriction is compatible with the root. It follows that the root is big, as thanks to Lemma 23 we may assume that $\varepsilon < 1/2$. Hence $S_0$ is just the set of good restrictions.
At depth $d$ in the tree, by the definition of compatibility each restriction in $S_d$ is compatible with exactly one big node. Consider any such big node $z$. It is labelled with a $k$-CNF $B$ and has a NO child $z_0$ and a YES child $z_1$. Define $P_z$ as

$$\Pr[\rho \text{ is not compatible with either } z_0 \text{ or } z_1 \mid \rho \text{ is compatible with } z]$$

This is equal to the probability that $\rho$ does not fix $B$, under the condition that $\rho$ is good and $\rho \in C_z$, where $C_z$ is the set of restrictions which correctly fix all CNFs above $z$. Notice that $\Pr[\rho \text{ is bad}] = \varepsilon$, so in particular $\Pr[\rho \text{ is bad}]) \leq \varepsilon$. On the other hand $\Pr[\rho \text{ is good} \implies \rho \text{ is bad}] = \varepsilon$, since $z$ is big.

Hence, summing over big nodes at level $d$, the probability that $\rho$ is compatible with some (not necessarily big) node at depth $d+1$ is at least

$$\sum_{z \in \{0, 1\}^d, z \text{ big}} (1 - P_z) \Pr[\rho \text{ is compatible with } z] \geq (1 - 12kn^{-1/7}) \Pr[S_d].$$

To obtain $S_{d+1}$ we must finally remove the restrictions which are compatible with non-big nodes at depth $d+1$. But there are at most $2^{d+1}$ such nodes, so the probability of being compatible with any of them is at most $2^{d+1} \varepsilon$. A straightforward calculation shows that

$$(1 - 12kn^{-1/7})(1 - d \cdot 12kn^{-1/7} - 2^{d+1} \varepsilon) - 2^{d+1} \varepsilon \geq 1 - (d+1) \cdot 12kn^{-1/7} - 2^{d+2} \varepsilon,$$

which completes the inductive step.

## 5 Non-reducibility

Consider a restriction $\rho$ and a $\Sigma_1^b$ formula $\exists y < t \theta(a, y)$, where $\theta$ is a PV formula and $a$ is some number. We say that this formula is witnessed in $\rho$ if there is some $b < t$ such that $\theta(a, b)$ holds in $\rho$. That is, if you run the computation verifying $\theta(a, b)$ and answer queries to $\alpha$ with values from $\rho$, these values are all defined and the computation is accepting.

Recall from Definition 6 that a precomputation of a $P_{NP}$ machine contains a correct witness for every YES reply, but may be wrong about NO replies.

**Definition 28.** Let $\rho$ be a restriction. A precomputation $w$ of a $P_{NP}$ machine $M$ is fixed by $\rho$ if both of the following hold.

1. For every NP query in $w$ with a YES reply, the witness provided by $w$ is correct in $\rho$.
2. No NP query in $w$ with a NO reply is witnessed in any legal $\sigma \supseteq \rho$.

We say that $\rho$ fixes a precomputation of $M$ on input $v$ if there is some such $w$. For a function $f$ computed by a $P_{NP}$ machine, we write $\rho \models w : f(x) = y$ if $\rho$ fixes a precomputation $w$ of $f$ on input $x$ that outputs $y$, and we write $\rho \models f(x) = y$ if $\rho$ fixes some such $w$. 

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If \( w \) is fixed by \( \rho \) then \( \rho \) fixes, in the sense of Definition 24, each DNF representing an NP query made in \( w \). Note that \( w \) does not have to be a computation of \( M \) relative to any complete oracle \( \alpha \). In fact, if \( \rho \) is legal and \( w \) contains a query like “is there a witness to CPLS?”, then \( w \) cannot be a computation of \( M \) relative to any complete \( \alpha \), because \( \alpha \) will contain a witness to a YES answer, but the answer given in \( w \) will be NO.

The symbol \( \vdash \) is intentionally chosen to be the same one as in forcing. In fact, one could formulate the concept of fixing in terms of a forcing relation, with the restrictions as forcing conditions. However, attempting to preserve all the trappings of forcing in the context of finite combinatorics leads to some annoying issues, so in this paper we do not explore this possibility further.

**Lemma 29.** For a \( P^{NP} \) function \( f \), a restriction \( \rho \) and an input \( x \), there is at most one \( y \) such that \( \rho \vdash f(x) = y \).

**Proof.** The progress of a \( P^{NP} \) precomputation depends only on the YES/NO replies to NP queries, not on the witnesses chosen. In all precomputations of \( f(x) \) fixed by \( \rho \) these replies are necessarily the same. \( \square \)

Below a “suitable” \( n \) is one for which \( n^{1/7} \) is a power of two.

**Lemma 30.** Let \( M \) be a \( P^{NP} \) machine, running on inputs \( x \) with \( |x| \) polylogarithmic in \( n \). For all suitable large enough \( n \), for every such input \( x \),

\[
\Pr_{\rho \sim R_{p,q}}[\rho \text{ fixes a precomputation of } M \text{ on } x] \geq 1 - n^{-1/6}.
\]

**Proof.** We can model a run of \( M \) on \( v \) as a decision tree \( T_M \). The height \( d \) of \( T_M \) is bounded by the running time of \( M \). At each node the tree makes an NP query; by negating the reply, we can view this as a query to a \( k \)-CNF, where \( k \) is some obvious syntactic upper bound on the time needed to verify a witness to the query. Since \( M \) is a \( P^{NP} \) machine, \( k \) can be chosen polynomial in the running time of \( M \). So we can apply Lemma 27 with \( k = d = |n|^c \) for some \( c \in \mathbb{N} \). This gives the lower bound

\[
1 - |n|^{2c}n^{-1/7} - 2^{|n|^{c}+1} \varepsilon
\]

on the probability that \( \rho \) is compatible with one of the leaves of \( T_M \). By Lemma 29, the probability \( \varepsilon \) that \( \rho \) is bad is exponentially small in \( n \), so the bound in (6) is at least \( 1 - n^{-1/6} \) for \( n \) sufficiently large.

Finally, suppose \( \rho \) is compatible with a leaf of \( T_M \). We form a precomputation \( w \) by answering queries with the replies given on the path from the root to the leaf. For each YES reply, it follows from the definition of fixing a DNF to 1 (that is, fixing a CNF to 0) that \( \rho \) provides enough information to verify at least one witness to the reply; we make some such witness part of \( w \). \( \square \)

**Lemma 31.** Let a search problem in \( rWPHP_2 \) be given by \( P^{NP} \) functions \( f_s(u) \) and \( g_s(v) \). Let \( s(n) \) be quasipolynomial in \( n \). Then for all suitable large enough \( n \),

\[
\Pr_{\rho \sim R_{p,q}}[\text{there exist } v \neq v' \text{ such that } \rho \vdash f_s(g_s(v)) = v'] \geq 1 - 3n^{-1/6}.
\]
\textit{Proof.} Choose $n$ sufficiently large. Let $M$ be the \textsf{P}^{\textsf{NP}} machine which takes input $n, v$ and computes $f_s(g_s(v))$ by first computing $g$ and then $f$. As $s$ is quasipolynomial in $n$, we may assume that $M$ satisfies the assumption of Lemma 30 on input size. We will write just $f$ and $g$ below, suppressing the parameter $s$.

Consider the machine $M$ running on inputs $v < 2s$. By Lemma 30, for any fixed $v$, a random $\rho$ fixes a precomputation of $M$ on $v$ with probability at least $1 - n^{-1/6}$. It follows that with probability at least $1 - 3n^{-1/6}$ a random $\rho$ simultaneously fixes precomputations for at least $2/3$ of all inputs $v$ (as otherwise the fraction of pairs $(\rho, v)$ in which $\rho$ does not fix a precomputation on $v$ would be more than $n^{-1/6}$). Fix such a $\rho$. In particular, there are at least $s + 1$ many distinct inputs $v_0, \ldots, v_s$ for which $\rho$ fixes precomputations $w_0, \ldots, w_s$.

The machine $M$ first computes $u = g(v)$, which is necessarily less than $s$, and then computes $f(u)$. Hence, by the pigeonhole principle, there is some $u$ for which there exist distinct $i, j$ such that $\rho \models w_i : g(v_i) = u$ and $\rho \models w_j : g(v_j) = u$. (Here and below we are abusing our notation slightly, as $w_i$ and $w_j$ are really precomputations of $g$ followed by $f$.)

But, by Lemma 29 there must be a single $v'$ such that $\rho \models w_i : f(u) = v'$ and $\rho \models w_j : f(u) = v'$. At least one of $v_i$ and $v_j$ is distinct from $v'$; without loss of generality suppose $v_i$ is. Let $v = v_i$ and $w = w_i$. Thus we have $\rho \models w : f(g(v)) = v'$ and $v' \neq v$, as required. \hfill \square

Now consider the following Prover-Adversary game, given by an \textsf{NP} search problem $Q(x, y)$ and a $\Sigma_2^p$ search problem $R(x', y')$. At the start of the game, the Prover queries $R$ for some input $x'$, with $|x'|$ polynomial in $|x|$, and the Adversary gives a reply $y'$. Then the Prover repeatedly queries bits of the oracle $\alpha$, and the Adversary replies. The Prover is limited in the number of bits of $\alpha$ he can remember at once, and can also forget bits to save memory. The Prover wins when the partial oracle in his memory either witnesses $Q(x, y)$ for some $y$, or witnesses that $R(x', y')$ is false. This game models \textsf{PLS} counterexample reducibility, in the following sense.

\textbf{Lemma 32.} Suppose an \textsf{NP} search problem $Q(x, y)$ is cleanly \textsf{PLS} counterexample reducible to a $\Sigma_2^p$ search problem $R(x', y')$. Then for all inputs $x$ the Prover can win the game using only polynomially many (in $|x|$) bits of memory.

\textit{Proof.} This is an immediate consequence of the definitions of \textsf{PLS} counterexample reducibility (Definition 3) and \textsf{PLS}. Using the notation of Definition 3, the Prover first asks for $y'$ such that $R(e(x), y')$. Since the reduction is clean, the value of $e(x)$ does not depend on any oracle queries (Definition 17). He then sets $x'' = (x, y')$ and simulates the (exponential time, but polynomial memory) task of solving the \textsf{PLS} problem $P(x'', y'')$ by starting with $y'' = 0$ and then repeatedly setting $y''$ to $N_{x''}(y'')$, finding domain elements of smaller and smaller cost, until either the costs stop decreasing or $y''$ leaves the domain $F_{x''}$. Once $y''$ satisfying $P(x'', y'')$ is found, the Prover simulates the polynomial-time computation of $y := d(x, y'')$. Then he simulates two additional polynomial-time computations in order to check whether $Q(x, y)$ holds and whether $y$ witnesses
that \( R(x', y') \) is false. By the definition of PLS counterexample reducibility, one of the two possibilities must hold, thus allowing the Prover to win the game.

During the entire game, the Prover never needs to remember more bits of the oracle than are necessary to fix simultaneously the cost and membership of the domain of one candidate solution to \( P \), the computation of its neighbour, the cost of the neighbour, and possibly computations of \( d \) and the witnessing for \( Q \) and \( R \).

We can now prove Theorem 13 that CPLS is not in the class APPROX.

**Proof of Theorem 13.** Assume that CPLS is in APPROX. Then by Corollary 21 it is cleanly PLS counterexample reducible to an instance of rWPHP2 given by some functions \( f \) and \( g \). This means that, by Lemma 22, the Prover can win the Prover-Adversary game in which \( Q \) is CPLS and \( R \) is rWPHP2, using only polynomially many bits of memory. We obtain a contradiction by describing a strategy for the Adversary that defeats any Prover with small memory.

The Prover first makes his query \( x' \) to rWPHP2. The Adversary then picks a restriction \( \rho \) from \( \mathcal{R}_{n,q} \) for which there exist a precomputation \( w \) and numbers \( v, v' < 2x' \) with \( v \neq v' \) such that \( \rho \vdash w : f_{x'}(g_{x'}(v)) = v' \). By Lemma 21 such a \( \rho \) exists, and by Lemma 23 we may further assume that it is good. The Adversary replies with \( \langle v, w \rangle \).

Then, using the goodness of \( \rho \) and the limited size of the Prover’s memory, the Adversary is able to have in hand throughout the game a legal \( \sigma \supseteq \rho \) which contains all bits in the Prover’s current memory. Such a \( \sigma \) can never witness CPLS, because it is legal. However, a legal \( \sigma \) also cannot witness that \( \langle v, w \rangle \) is not a solution to rWPHP2, because the only way to do this would be to witness that one of the NO replies in \( w \) is wrong, which is impossible by the choice of \( \rho \). The details of the strategy are as in the proof of [42, Theorem 5.10].

### 6 Reformulation in propositional logic

In this section we sketch another way of presenting our main result about bounded arithmetic, that CPLS, considered as a \( \forall \Sigma_{b1}^p \) principle, is not provable in \( \text{APC}_2 \). We will use propositional proof complexity and in particular the well-known Paris-Wilkie translation of relativized bounded arithmetic into propositional logic [37].

Suppose \( \varphi \) is bounded formula of \( L_{PV} \), and that we have specified values \( \bar{n} \) for all free variables in \( \varphi \). We can write a propositional formula \( \langle \varphi \rangle \) with the same semantics as \( \varphi \), if we interpret propositional variables \( x_n \) as bits \( \alpha(n) \) of the oracle. Below we will use *narrow* to mean “of width polylogarithmic in \( \bar{n} \)”.

If \( \varphi \) does not mention the oracle \( \alpha \), then its translation \( \langle \varphi \rangle \) is the propositional constant \( \top \) or \( \bot \), depending on whether \( \varphi \) is true or false in \( \mathbb{N} \). If \( \varphi \) is \( \alpha(n) \), then \( \langle \varphi \rangle \) is the propositional variable \( x_n \). If \( \varphi \) is a PV formula, then \( \langle \varphi \rangle \) is a narrow CNF — we can take it to be the conjunction of clauses expressing “some oracle reply in \( w \) is false” over all possible rejecting computations \( w \).
of the polynomial-time machine deciding $\varphi$. If $\varphi$ is a $\Pi_b^1$ formula $\forall x < n \theta(x)$, then again $\langle \varphi \rangle$ is a narrow CNF, namely the conjunction, over $m < n$, of the translations $\langle \theta(x) \rangle$ with $x \mapsto m$.

The translation theorem we will use follows from the translation of $T_2^1$ into treelike $\text{Res}(\log)$ refutations from [26] and the connection between treelike $\text{Res}(\log)$ and narrow resolution [29]. It can also be shown via PLS witnessing, as described in [11].

**Theorem 33.** Let $\varphi(n)$ be a $\Pi_b^1$ formula and suppose $T_2^1 \vdash \forall n \neg \varphi(n)$. Then the translations $\langle \varphi \rangle$ have narrow resolution refutations.

Now suppose for a contradiction that $\text{APC}_2 \vdash \text{CPLS}$. Consider CPLS as described in Section 4, with parameters $a = b = n$ and $c = \lfloor n^{1/7} \rfloor$ and the structure of the problem given entirely by the oracle. Let $Q(n, y)$ assert that $y$ is a solution to CPLS on input $n$. We may bound $y$ by some term $t(n)$, such that $\text{APC}_2 \vdash \forall n \exists y < t Q(n, y)$. By the proof of Lemma 18, there exist $\text{P}^{\text{NP}}$ machines $f$, $g$ defining an instance of $r\text{WPHP}_2$, and a term $s(n)$, such that

$$T_2^1 \vdash \forall n \forall v < 2s \forall w \left[ w \text{ is not a computation of } f(g_s(v)) \lor \text{output}(w) = v \lor \exists y < t Q(n, y) \right].$$

Let $M$ be the $\text{P}^{\text{NP}}$ machine which takes input $n, v$ and computes $f(g_s(v))$ by first computing $g$ and then $f$. We think of $v$ as the “real input” to $M$ and of $n$ as a parameter, and write $\text{Comp}_{M}(v, w)$ for the $\Pi_b^1$ formula from Definition 6 expressing that $w$ is a computation of $M$ on input $v$. Noting that the expression on the right above is $\forall \Sigma_b^1$, we can apply Theorem 33 to conclude that the family of narrow CNFs

$$\Phi_{n,v,w} := \langle v < 2s \rangle \land \langle \text{Comp}_{M}(v, w) \rangle \land \langle \text{output}(w) \neq v \rangle \land \bigwedge_{y < t} \langle \neg Q(n, y) \rangle$$

has narrow resolution refutations, that is, of width polylogarithmic in $n, v, w$.

Fix a suitable large $n$. By definition, no legal restriction $\sigma$ can falsify any clause in the last conjunct $\bigwedge_{y < t} \langle \neg Q(n, y) \rangle$, as otherwise for some $y$ there is an accepting computation of $Q(n, y)$ over $\sigma$, so $\sigma$ witnesses CPLS.

By Lemma 31 with high probability for a random $\rho$ from $\mathcal{R}_{p,q}$ there exist $v < 2s$ and a precomputation $w$ of $M$ on $s$ with output$(w) \neq v$ such that $w$ is fixed by $\rho$, meaning that all witnesses in $w$ to YES answers are correct in $\rho$ and no query with a NO answer has a witness in any legal extension of $\rho$. It follows that for such $v, w$, no clause in the first three conjuncts of $\Phi_{n,v,w}$ is false in any legal extension of $\rho$. By Lemma 23 we can pick a good $\rho$ for which such $v, w$ exist.

By the Prover-Adversary construction in the proof of [42, Theorem 5.10], we can exploit the limited width of the refutation of $\Phi_{n,v,w}$ to find a legal extension of $\rho$ which falsifies one of the conjuncts of $\Phi_{n,v,w}$. This is a contradiction.
7 Open problems

The random resolution propositional proof system was introduced in [11]. Very roughly speaking, a refutation of a CNF $F$ in this system is a refutation of $F \land A$, where $A$ is any CNF which is true with high probability.

Suppose a sentence $\forall n \varphi(n)$, with $\varphi$ a $\Sigma_1^b$ formula, is provable in the sub-theory of APC$_2$ consisting of $T_2^2$ together with the surjective WPHP only for polynomial time functions. It was shown in [11] that this implies that the translations $\langle \lnot \varphi(n) \rangle$ have narrow refutations in random resolution.

Open Problem 1. Is there a natural propositional proof system which captures, in a similar way, the $\forall \Sigma_1^b$ consequences of full APC$_2$?

Ideally, one would want to show not only that APC$_2$ proofs translate into the system, but also something in the opposite direction, for example, that if $\langle \lnot \varphi(n) \rangle$ has small, suitably uniform refutations in the system, then $\forall n \varphi(n)$ is provable in APC$_2$. Some system with these properties could be constructed using the Paris-Wilkie translation and our arguments in Section 6, but it would be rather unnatural and awkward.

It is consistent with what we know that narrow random resolution, or possibly random resolution with no width restriction, already provides a positive answer to Open Problem 1. So, we can ask:

Open Problem 2. Is there a $\forall \Sigma_1^b$ sentence which is provable in APC$_2$ but whose propositional translations do not have narrow random resolution refutations?

A candidate is the Herbrandized ordering principle HOP, which is provable in APC$_2$ [11] but not in the subtheory mentioned above [2].

What makes this problem interesting is that, so far, our only tool for proving lower bounds on random resolution is the fixing lemma of [12]. For a typical random restriction, it is a small step from proving this to proving our conditional fixing lemma from Section 4, which implies unprovability in APC$_2$. But showing a separation seems to require finding a principle and a random restriction for which one lemma holds, but not the other. The restrictions used to show unprovability of HOP in [2] may be useful here.

Finally we mention a rather obvious question: is every problem in APPROX reducible to CPLS? This is subsumed in the old open problem, discussed in the introduction, of separating the classes GI$_k$ or the theories $T_k^2$: it is possible that every search problem reducible to any GI$_k$ is already reducible to CPLS.

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