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Lifting to Bounded-Depth and Regular Resolutions over Parities via Games*

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Abstract

Proving superpolynomial lower bounds on proof size in the proof system resolution over parities (Res($\oplus$)) remains a significant open challenge. A recent breakthrough by Efremenko, Garlik, and Itsykson (STOC 2024) established an exponential lower bound for regular Res($\oplus$).

In this work, we introduce a lifting technique for regular Res($\oplus$), applicable to a wide range of formulas. Specifically, we develop a method that transforms any formula with large resolution depth into a formula requiring exponential-size regular Res($\oplus$) refutations. This transformation is achieved through a combination of mixing and constant-size lifting.

Using this approach, we provide an alternative and improved separation between resolution and regular Res($\oplus$), originally proved by Bhattacharya, Chattopadhyay, and Dvorak (CCC 2024). We construct an n-variable formula with a polynomial-size resolution refutation of depth $O(\sqrt{n})$, yet requires regular Res($\oplus$) refutations of size $2^{\Omega(\sqrt{n})}$.

Furthermore, we apply our technique to establish an exponential lower bound on the size of depth-$cn \log\log n$ Res($\oplus$) refutations, where n is the number of variables in the refuted formula, and c is a constant. The hard instances in this setting are Tseitin formulas lifted with the Maj_5 gadget. Since even depth-n Res($\oplus$) captures all possible definitions of regular Res($\oplus$), our result yields an exponential lower bound for top-regular Res($\oplus$), resolving an open question posed by Gryaznov, Pudlák, and Talebanfard (CCC 2022).

1 Introduction

Propositional proof systems are used to certify that given Boolean formulas are unsatisfiable. Cook and Rekhow [11] noticed that NP $\neq$ coNP implies that for every propositional proof system, there is a family of hard formulas that require superpolynomial proof sizes. However, currently, we cannot prove superpolynomial proof-size lower bounds for many particular proof systems.

One of the long-standing open questions in proof complexity is proving superpolynomial lower bounds on the size of derivations in Frege proof systems. Proof lines in Frege systems are Boolean

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formulas; each particular Frege system is defined by a sound and implicationally complete set of rules. Currently, we only know how to prove Frege lower bounds in bounded-depth cases where proof lines are restricted to be bounded-depth formulas over $\neg, \vee$ and $\wedge$ (see [1], for example). The techniques used to prove those lower bounds are quite similar to techniques used in bounded-depth circuits lower bounds. So, it was conjectured that techniques used by Razborov and Smolenski [27, 28] to prove a lower bound for constant depth circuits built up from $\neg$, $\vee$, $\wedge$, and Mod_p gates can be extended to bounded-depth Frege operating with formulas using $\neg$, $\vee$, $\wedge$ and Mod_p gates (denoted AC^0[p]-Frege). However, proving lower bounds for AC^0[p]-Frege is still open for all values of $p > 1$.

The weakest subsystem of AC^0[2]-Frege for which we still do not know superpolynomial lower bounds is resolution over parities (Res($\oplus$)). The proof lines in this proof system are disjunctions of linear equations over $\mathbb{F}_2$ (or, equivalently, negations of $\mathbb{F}_2$-linear systems), called linear clauses. A Res($\oplus$) refutation of an unsatisfiable CNF formula φ is a sequence of linear clauses $C_1, C_2, \ldots, C_s$ such that (1) C_s is the empty clause (i.e. identically false); (2) for every i, C_i is either a clause of φ or is obtained from C_j and C_k with $j, k < i$ by the resolution rule, or is obtained from C_j with $j < i$ by the weakening rule. The resolution rule allows to derive the clause $C \vee D$ from clauses $C \vee (f = 0)$ and $D \vee (f = 1)$, where f is a linear form. The weakening rule allows us to derive D from C if C semantically implies D. Recently, this proof system received a lot of attention from different researchers. Below, we highlight some of the achievements.

Tree-like lower bounds. There are plenty of tree-like Res($\oplus$) lower bounds for particular formulas obtained by different techniques: Prover-Delayer games [22, 23, 17, 18], reductions from communication complexity [22, 23, 21, 24], reductions from polynomial calculus degree [15].

Chattopadhyay et al. [9] proved that resolution depth can be lifted by stifling gadgets to tree-like Res($\oplus$) size. Independently, Beame and Koroth [6] got similar results. Resolution depth of an usatisfiable CNF formula φ equals the query complexity of finding a clause of φ that is falsified by the given assignment. Proving lower bound on resolution depth is relatively easy; so this lifting theorem gives us a relatively easy way to prove tree-like Res($\oplus$) lower bounds for many formulas.

Regular Res($\oplus$) lower bounds Recently, Efremenko, Garlik and Itsykson [13] proved the first exponential lower bounds on the size of regular (bottom-regular) Res($\oplus$) refutations. Regular Res($\oplus$) is a fragment of Res($\oplus$) in which resolving linear clauses C_1 and C_2 on a linear form f is permitted only if, for both $i \in \{1, 2\}$, the linear form f does not lie within the linear span of all linear forms that were used in resolution rules during the derivation of C_i. Regular Res($\oplus$) is known to be strictly stronger than tree-like Res($\oplus$).

A formula that was shown by [13] to be hard for regular Res($\oplus$) is the Binary Pigeonhole Principle (BPHP). The key technical tool for proving this lower bound is the notion of *closure*; given a $\mathbb{F}_2$-linear system in variables of BPHP, the closure is roughly speaking the set of pigeons on which this linear system is actually talking about.

Bhattacharya, Chattopadhyay, and Dvorák [7] have recently shown that specific CNF formulas require an exponential size refutation in regular Res($\oplus$) but admit polynomial size refutation in resolution. This result heavily utilizes the techniques from [13] and the techniques from lifting literature. However, unlike [9], this result is formula-specific and does not immediately provide a complexity measure that can be lifted to regular Res($\oplus$) size. The possibility of general lifting was left as an open question.

Other definitions of regularity. The question of what is the natural notion of regularity is debatable. For example, Gryaznov, Pudlak and Talebanfard [19] suggested the notion of *top-regular* Res($\oplus$) refutations, where if a linear clause C is derived by resolving over a linear form f, then f does not belong to the span of all linear forms that were used in resolution rules in all clauses that were derived using C. Gryaznov Pudlak and Talebanfard [19] posed an open question to prove superpolynomial lower bounds for top-regular Res($\oplus$) refutations. One can also define *weakly regular* Res($\oplus$), where for each path from the axioms to the contradiction, all resolved linear forms are linearly independent. Both top-regular and bottom-regular refutations are weakly regular. We also notice that the depth of any weakly regular refutation does not exceed the number of variables in the refuted formula. So, proving superpolynomial lower bounds on the size Res($\oplus$) refutations of depth at most n, where n is the number of variables in the refuted formula, will capture all possible definitions of regularity.

1.1 Our Contributions

Consider a CNF formula $\varphi(y_1, y_2, \ldots, y_m)$. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a gadget. Consider a lifted formula $\varphi \circ g$ that is obtained from φ by applying the substitutions $y_i := g(x_{i,1}, x_{i,2}, \ldots, x_{i,\ell})$ for all $i \in [m]$ and then converting the resulting formula in CNF.

We contribute to lifting with stifling gadgets introduced by Chattopadhyay et al. [9]. A gadget g is called k-stifling [9] if for every $a \in \{0,1\}$ and every $\ell - k$ variables of g, we can fix them such that regardless of the value of the rest k variables, the value of the gadget will be fixed to a. It is easy to see that the majority function $Maj_{2k+1} : \{0,1\}^{2k+1} \to \{0,1\}$ is k-stifling for every k.

To illustrate our approach, we present a clear and simple example of the lifting theorem.

Theorem 1.1 (Theorem 3.1)**.** *Let $g\colon \{0,1\}^\ell \to \{0,1\}$ be a 1-stifling gadget and φ be a CNF formula. Assume that $\varphi \circ g$ has a* Res($\oplus$) *refutation of rank W. Then φ has a resolution refutation of width at most W.*

Next, we contribute to the research direction initiated by Bhattacharya, Chattopadhyay, and Dvořák [7], which seeks a relatively simple proof-complexity measure for formulas that enables lifting lower bounds to the size of regular Res($\oplus$) refutations. Our result is not a classical lifting theorem, as it does not rely solely on composition with a gadget. Before applying composition, we first modify the formula using an operation called *mixing*. Importantly, mixing does not alter the formula's semantic meaning. Alekhnovich et al. [2] introduced a transformation of CNF formulas—also known as Alekhnovich's trick—where each clause C is replaced by $C \vee x$ and $C \vee \neg x$, with x being a randomly chosen variable. This transformation was used by [2] to separate regular resolution from general resolution. Our notion of mixing follows a similar idea but extends it by adding multiple random variables (either a constant or logarithmic number) to each clause. Specifically, we transform a clause C into $\bigwedge_{\alpha \in \{0,1\}^k} C \vee x_1^{\alpha_1} \vee \ldots \vee x_k^{\alpha_k}$, where x^0 denotes $\neg x$ and x^1 denotes x.

Theorem 1.2 (Informal restatement of Theorem 6.8)**.** *Let φ be an unsatisfiable CNF formula in n variables that requires a resolution depth at least αn. Let* mix(φ) *denote the mixing of φ, where every clause of φ is "mixed" with $O(\frac{1}{\alpha^2})$ additional variables. Then any regular* Res($\oplus$) *refutation of* mix(φ) $\circ\, h$ *has size at least $2^{\alpha n/4 - 1}$, where $h = \oplus_5 \circ g$ is the composition of the parity function and a 2-stifling gadget g (for example, the 5-bit majority (Maj_5) is 2-stifling).*

Using ideas from Theorem 1.2, we give an alternative proof of separation between regular Res($\oplus$) and Resolution. Namely, we give the family of formulas with M variable that has a resolution refutation of size poly(M) and depth $O(\sqrt{M})$ but requires regular Res($\oplus$) refutations of size $2^{\Omega(M^{1/2})}$. This improves the size lower bound $2^{\Omega(M^{1/12})}$ obtained in [7].

Theorem 1.3 (Informal restatement of Corollary 7.3 and Theorem 7.5)**.** *Let G_n be a graph of $n \times n$ grid. Let* $\mathrm{mix}(\mathrm{Peb}(G_n))$ *denote some specific mixing of the pebbling formula based on G_n, where each clause is mixed with $O(\log n)$ variables. Then there exists a constant k and a gadget $g : \{0,1\}^k \to \{0,1\}$ such that* $\mathrm{mix}(\mathrm{Peb}(G_n)) \circ g$ *requires regular* Res($\oplus$) *refutations of size at least $2^{n/4}$ but has a resolution refutation of size* poly(n) *and depth $O(n)$.*

Theorem 1.3 in particular implies that depth-n Res($\oplus$) (where n is the number of variables in the refuted formula) is more powerful than regular Res($\oplus$). We apply the developed lifting technique to prove an exponential lower bound for Res($\oplus$) refutations with depth up to $cn \log\log n$, where n is the number of variables and c is a small constant.

Theorem 1.4 (Informal restatement of Corollary 8.3)**.** *Let* $\mathrm{T}(G, c)$ *be an unsatisfiable Tseitin formula based on a d-regular expander with n vertices. Then, any* Res($\oplus$) *refutation of* $\mathrm{T}(G, c) \circ Maj_5$ *has either size at least 2^n or depth at least $\Omega(nd \log d)$. In particular, if $d = \Theta(\log n)$, then* $\mathrm{T}(G, c) \circ Maj_5$ *is a formula with $m = 5dn/2$ variables and size* poly(m) *such that any* Res($\oplus$) *refutation of* $\mathrm{T}(G, c) \circ Maj_5$ *has either size at least $2^{\Omega(m/\log m)}$ or depth at least $\Omega(m \log\log m)$.*

The direct consequences of Theorem 1.4 are the following:

- Exponential refutation size lower bounds for weakly regular and, thus, top-regular Res($\oplus$) are established (see Fig. 1). This answers the open question raised in articles [19] and [7].
- Exponential lower bounds for top-regular Res($\oplus$) imply exponential lower bounds for weakly read-once linear branching programs computing search problems (see [19] for details).

1.2 Overview of Technique

Lifting of strategies in games. A standard approach to proving lifting theorems — from a complexity measure μ_1 in a weak proof system Π_1 to a complexity measure μ_2 in a stronger proof system Π_2 — relies on an explicit transformation: given a Π_2-proof π_2 of a lifted formula $\phi \circ g$, one constructs a corresponding Π_1-proof π_1 of the base formula ϕ, such that $\mu_2(\pi_2)$ can be bounded in terms of $\mu_1(\pi_1)$. In contrast, our approach circumvents this proof-level transformation entirely. Instead, we work with game characterizations of the complexity measures μ_1 and μ_2, and show how a winning strategy for the game associated with μ_1 can be transformed into a strategy for the game associated with μ_2, applied to the lifted formula.

A simple instance of this idea was introduced by Urquhart [30], who demonstrated that resolution depth lower bounds can be lifted to tree-like resolution size lower bounds using the $\oplus_2$ gadget. His method involved transforming an Adversary strategy in the Prover–Adversary game characterizing the resolution depth of a formula ϕ into a Delayer strategy in the Prover–Delayer game characterizing the tree-like resolution size of the lifted formula $\phi \circ \oplus_2$. We provide several more non-trivial applications of this transformation between game strategies.

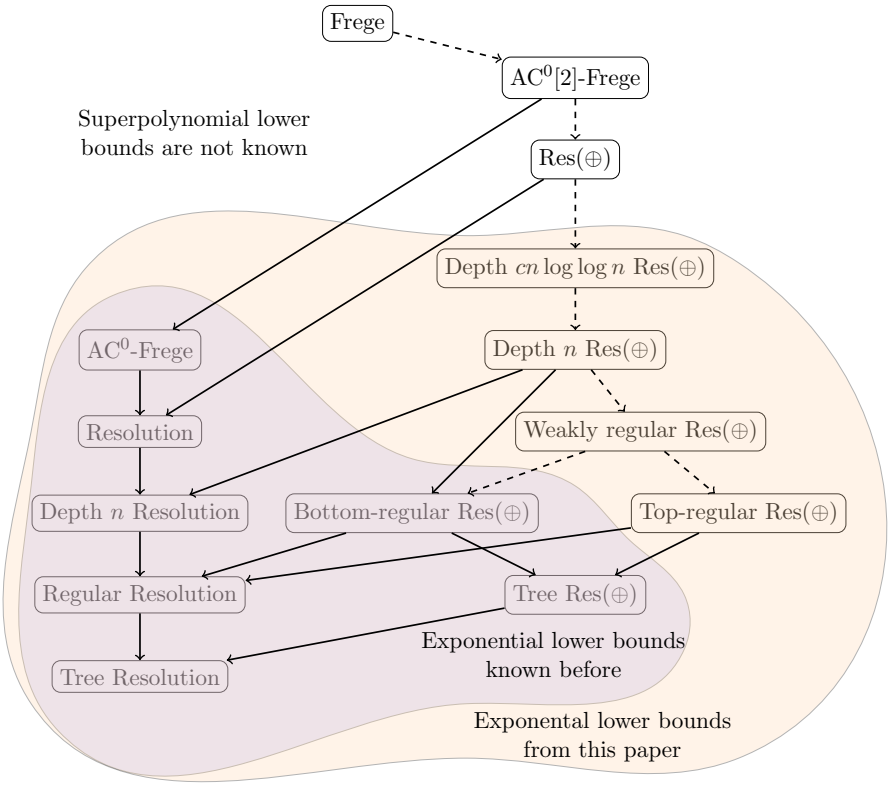

Figure 1: A summary of proof systems for which superpolynomial lower bounds are known. The arrows indicate p-simulations; the solid arrows indicate that it is known that there is no p-simulation in the other direction.

Closure in the lifting framework. Efremenko, Garlik, and Itsykson [13] introduced the notion of closure in the context of the binary pigeonhole principle. Subsequently, Bhattacharya, Chattopadhyay, and Dvorák [7] observed that this concept can also be applied in lifting.

Assume we are working within the lifting framework. Let $Y = \{y_1, y_2, \ldots, y_m\}$ denote the set of *unlifted* variables, and let $X = \{x_{i,j} \mid i \in [m], j \in [\ell]\}$ denote the corresponding set of *lifted* variables. These variables are connected via a *1-stifling* gadget g, in the following sense: for each $i \in [m]$, the values of the lifted variables $x_{i,1}, \ldots, x_{i,\ell}$ determine the value of the unlifted variable y_i through the function g, that is, $y_i = g(x_{i,1}, \ldots, x_{i,\ell})$.

Consider a $\mathbb{F}_2$-matrix A whose columns correspond to lifted variables. We say that a matrix A is *safe*[1] [13] if the linear span of the columns of A has a basis consisting of columns of A such that for every unlifted variable y_i, there is at most one basis element, corresponding to a column $x_{i,j}$ for some j. To solve a linear system with a safe matrix A, we can assign arbitrary values to

[1]Here, we give slightly simplified notions of safe matrices and closure of the matrices. In Section 2.4, we define the notion of a safe set of linear forms over lifted variables; there is the following correspondence: the set of linear forms is safe if its coefficient matrix is safe. In Section 2.5, we define the closure of the set of linear forms over lifted variables, and again, the closure of the set of linear forms equals the closure of its coefficient matrix.

all non-basis variables; the values of the basis variables are then uniquely determined to satisfy the system. Since the gadget g is 1-stifling, this implies that for any satisfiable linear system $Ax = b$ with safe A, and for every assignment σ to the unlifted variables, there exists a solution to $Ax = b$ that induces σ on the unlifted variables.

Now assume that an $\mathbb{F}_2$-matrix A, whose columns correspond to lifted variables, is not necessarily safe. A *closure* of A [13] is defined as an inclusion-minimal set of unlifted variables $I \subseteq Y$ such that removing from A all columns corresponding to I yields a safe matrix. It is known that the closure of A is unique and that its size does not exceed the rank of A. By the previously discussed property of safe matrices, it follows that for any solution π of a linear system $Ax = b$, and for any full assignment σ to the unlifted variables such that π agrees with σ on the closure of A, there exists a solution to $Ax = b$ that induces σ on all unlifted variables. Roughly speaking, the closure of a matrix is the set of unlifted variables that the matrix "talks about."

Using the notion of closure, we define a correspondence between linear systems over lifted variables and partial assignments to unlifted variables as follows:

— A linear system $Ax = b$ is said to *correspond* to a partial assignment ρ if there exists a solution σ of $Ax = b$ such that ρ is the restriction to the closure of A of the assignment induced by σ on the unlifted variables.

This correspondence is not necessarily one-to-one. Crucially, if $Ax = b$ corresponds to a partial assignment ρ, and ρ does not falsify any clause of a CNF formula φ, then $Ax = b$ is consistent with every clause of the lifted formula $\varphi \circ g$.

1.2.1 Lifting from resolution width to Res(⊕) rank

The proof of Theorem 1.1 relies on *Spoiler-Duplicator* games introduced by Atserias and Dalmau [5], which characterize resolution width, as well as on analogous games characterizing Res(⊕) rank (or width) [18] (see Section 3 for the formal definitions of these games). In both settings, Duplicator's strategies are defined in terms of sets of winning positions: partial assignments in the case of resolution, and linear systems in the case of Res(⊕).

Assuming that φ does not admit a resolution refutation of width at most W, there exists a winning strategy for Duplicator in the $(W + 1)$-pebble resolution game. We construct from this a corresponding winning strategy for Duplicator in the $(W + 1)$-pebble Res(⊕)-game for the lifted formula $\varphi \circ g$. This implies that $\varphi \circ g$ does not have a Res(⊕) refutation of rank W.

The key to this transformation is the correspondence defined above: a linear system over lifted variables is considered a winning position in the Res(⊕) game if it corresponds to a winning position (i.e., a non-falsifying partial assignment) in the original resolution game (see Section 3 for details).

1.2.2 Prover Delayer games and random walk

Let φ be a CNF formula and $\mathcal{A}$ be a set consisting of partial assignments for variables of φ. We assume that $\mathcal{A}$ has two properties:

- $\mathcal{A}$ is closed under restrictions: if for every $\rho \in \mathcal{A}$ for every $\sigma \subseteq \rho$, $\sigma \in \mathcal{A}$.
- For every $\sigma \in \mathcal{A}$, σ does not falsify any clause of φ.

We define a $(\varphi, \mathcal{A})$-*game of Prover and Delayer with starting position* $\rho_0 \in \mathcal{A}$. In this game, two players, Prover and Delayer, maintain a partial assignment ρ for variables of φ that initially equals ρ_0. On every move, Prover chooses a variable x, and Delayer has two options:

- Delayer can earn a *white* coin and reports $*$. Then, Prover chooses a Boolean value a of x.
- Delayer can earn a *white* coin and pay a *black* coin to choose a Boolean value a of x by himself.

The current assignment ρ is updated: $\rho := \rho \cup \{x := a\}$. The game ends when $\rho \notin \mathcal{A}$.

Delayer's strategy is called *linearly described* if there exists a map f that takes as input an ordered set of variables L and a variable x, and returns either $*$ or an $\mathbb{F}_2$-affine function h depending on the variables in L. The strategy is applied as follows: given a game history $x_1 = a_1, x_2 = a_2, \ldots, x_k = a_k$ and a requested variable x, Delayer evaluates $f((x_1, x_2, \ldots, x_k), x)$. If $f((x_1, x_2, \ldots, x_k), x) = *$, then Delayer reports $*$. Otherwise, if $f((x_1, x_2, \ldots, x_k), x) = h$ for some affine function h, Delayer reports $h(a_1, a_2, \ldots, a_k)$.

Random walk. Our central technical tool for proving lower bounds on the size of regular and bounded-depth Res$\oplus$ proofs is a random walk on the refutation graph.

Let Π be a Res$(\oplus)$ refutation, C_0 a linear clause from Π, Σ a set of full assignments that falsify C_0, and $t \in \mathbb{N}$ a natural number. A (Π, C_0, Σ, t)-*random walk* is defined as follows: sample an assignment σ uniformly at random from Σ, and perform a random walk of length t on the refutation graph of Π, starting at the node labeled by C_0. At each step, the walk proceeds from a linear clause to a premise that is falsified by σ. If the walk terminates at a node labeled with a linear clause C, then C is the value of the random variable defined by the walk.

We prove the following theorem.

Theorem 1.5 (Simplified version of Theorem 4.3)**.** *Assume that in a* $(\varphi, \mathcal{A})$*-game with starting position* $\rho_0 \in \mathcal{A}$*, Delayer has a linearly described strategy that guarantees him to earn* w *white coins while paying at most* c *black coins. Let* $g : \{0,1\}^\ell \to \{0,1\}$ *be a 2-stifling gadget. Consider a* Res$(\oplus)$ *refutation* Π *of* $\varphi \circ g$*. Let* C_0 *be a linear clause from this refutation* Res$(\oplus)$ *refutation of* $\varphi \circ g$*. Assume that the linear system* $\neg C_0$ *corresponds to the partial assignment* ρ_0*. Let* Σ *be the set of all assignments* π *falsifying* C_0 *such that* π *defines the correspondence between* $\neg C_0$ *and* ρ_0*. (If* C_0 *is an empty clause, then* ρ_0 *is an empty assignment; hence, in this case,* Σ *is the set of all assignments.) Let a linear clause* C *denotes the result of the* (Π, C_0, Σ, t)*-random walk. Then with probability at least* $2^{-c(\ell-1)}$*, the linear system* $\neg C$ *corresponds to some* $\rho \in \mathcal{A}$*.*

Similar random walk arguments have appeared in the context of regular Res$(\oplus)$ lower bounds for specific formulas in [13] and [7]. In both works, the random walks are initiated from the empty clause. The analysis in [13] assumes a uniform distribution over assignments, whereas [7] employs a non-uniform distribution. Like [13], we adopt the uniform distribution; however, a key difference is that in our setting, the probability of reaching a good clause is exponentially small (albeit not negligible), in contrast to the constant success probabilities observed in [13] and [7]. While our proof shares some structural similarities with that of [13], our result is much more general, necessitating a more careful analysis and deeper arguments.

1.2.3 Lifting to Regular Resolution Size

We define several games on Boolean formulas and show that if a sufficiently strong strategy exists in a game based on a formula φ, then the lifted version of φ is hard for regular Res($\oplus$) proofs. We begin with the first game, which is particularly well-suited for the lifting argument. Subsequently, we introduce simpler games and demonstrate how strategies for these games can be lifted to strategies in the original game.

Let Φ be an unsatisfiable CNF formula represented as $\Phi = \bigwedge_{v \in V} \phi_v$, where each ϕ_v is itself a CNF formula in which all clauses share the same set of variables.

A partial assignment ρ is called *q-correct* for Φ if for every set $U \subseteq V$ such that $|\mathrm{Vars}(\bigwedge_{v \in U} \phi_v)| < |\mathrm{Vars}(\Phi)| - q$, ρ can be extended to an assignment satisfying $\bigwedge_{v \in U} \phi_v$, where $\mathrm{Vars}(\Phi)$ denotes the set of variables that appear in Φ.

Let $\mathcal{A}_q$ denote the set of all q-correct partial assignments for Φ.

Advanced (Φ, q)-games of Prover and Delayer. We define the advanced (Φ, q)-game of Prover and Delayer as the $(\Phi, \mathcal{A}_q)$-game of Prover and Delayer with *empty starting position.*

Theorem 1.6 (Theorem 5.2)**.** *Let Φ be an unsatisfiable CNF formula. Assume that Delayer has a linearly described strategy in the advanced (Φ, q)-game that guarantees him to earn t white coins while paying at most c black coins. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget. Then the size of any regular* Res($\oplus$) *refutation of $\Phi \circ g$ is at least $2^{t - q\ell - c(\ell-1)}$.*

The overall strategy for proving lower bounds on the size of regular Res($\oplus$) refutations follows the same approach as in [13] and later in [7]. We analyze a random walk within a regular Res($\oplus$) refutation of $\Phi \circ g$, guided by a random assignment σ. The walk begins at the empty clause and, at each step, proceeds to a premise falsified by σ. By Theorem 1.5, with noticeable probability, this walk reaches a linear clause C such that $\neg C$ corresponds to a q-correct partial assignment.

We then show that if the distance from C to the empty clause is large and $\neg C$ is q-correct, regularity forces C to contain many linearly independent linear forms. This, in turn, implies that the probability of σ falsifying such a clause C is very small. Therefore, the refutation must include many such complex linear clauses, leading to a lower bound on its size.

This theorem enables us to resolve an open question posed in [7] and establish a lower bound for lifted Tseitin formulas:

Corollary 1.7 (Informal restatement of Corollary 5.4)**.** *Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget and G be is a good enough constant-degree expander on n vertices. Then the size of any regular* Res($\oplus$) *refutation of $\mathrm{T}(G, c) \circ g$ is at least $2^{\Omega(n)}$, where $\mathrm{T}(G, c)$ is an unsatisfiable Tseitin formula based on the graph G.*

To derive Corollary 1.7 from Theorem 1.6, we introduce a *natural* strategy for the Delayer on Tseitin formulas that preserves the following invariant: at every stage, the only unsatisfiable connected component of the current Tseitin formula is the largest one. See Section 4.2 for details.

Later, we will demonstrate that lifted Tseitin formulas are also hard for an even stronger proof system.

A disadvantage of Theorem 1.6 is that advanced (Φ, q)-games of Prover and Delayer are a bit complicated. Our goal is to demonstrate that strategies from much simpler games can be lifted to strategies in the advanced Prover-Delayer games. We achieve this simplification in two stages, ultimately reducing the problem to strategies in very simple games that characterize resolution depth (see Section 6.2 for a formal definition of these games).

Simplified (Φ, q)-games of Prover and Delayer. Let us define simplified (Φ, q)-games , played between two players: Prover and Adversary. In each round, Prover selects a variable x from the formula Φ, and Adversary responds by assigning it a value of 0 or 1. The game continues until the current partial assignment ceases to be q-correct. For every move made, Adversary earns one coin.

Lemma 1.8 (Lemma 6.1)**.** *Assume that there is a strategy of Adversary in the simplified (Φ, q)-game that allows him to earn at least t coins. Let $\oplus_r : \{0,1\}^r \to \{0,1\}$ be the parity function. Then for the advanced $(\Phi \circ \oplus_r, qr)$-game, there is a linearly described strategy of Delayer that guarantees him to earn tr white coins while paying at most t black coins.*

The proof of Lemma 1.8 builds on the same idea that Urquhart used to lift resolution depth to tree-like resolution size in [30].

Lifting from resolution depth to regular Res($\oplus$) **size.** Plan of the proof of Theorem 1.2 is the following:

1. We lift a strategy of Adversary in the depth-characterizing game for formula ϕ allowing to earn him αn coins to a strategy of Adversary in the simplified $(\mathrm{mix}(\varphi), \epsilon n)$-game allowing him to earn $\alpha n/2$ coins.

2. We lift the latter strategy to the linearly described strategy of Delayer in the advanced $(\mathrm{mix}(\varphi) \circ \oplus_5, \epsilon n)$-game by Lemma 1.8.

3. Get lower bound on size of regular Res($\oplus$) refutation of $\mathrm{mix}(\varphi) \circ \oplus_5 \circ Maj_5$ refutations by Theorem 1.6 using that Maj_5 is a 2-stifling gadget.

By Corollary 1.7, expander-based Tseitin formulas lifted by a constant-size gadget are $O(1)$-CNF formulas with n variables of size $O(n)$ which require regular Res($\oplus$) refutation of size $2^{\Omega(n)}$. This is the best possible lower bound up to a constant in the exponent, and such tight lower bounds for regular Res($\oplus$) were not known before. Theorem 1.2 allows us to construct many formulas given the same tight lower bounds. To do this, we can apply Theorem 1.2 to $O(1)$-CNF formulas with n variables, $O(n)$ clauses and with resolution depth $\Omega(n)$. It is well-known that resolution depth is at least resolution width (see, for example, [30]). Hence, we can, for example, apply Theorem 1.2 to random $O(1)$-CNF formulas with n variables and $O(n)$ clauses that are known to have resolution width $\Omega(n)$ [3].

1.2.4 Size vs Depth Tradeoff

Using the techniques developed in this paper, we obtain the following result.

Theorem 1.9 (Theorem 8.1)**.** *Assume that there are integers t and c such that for every $\rho \in \mathcal{A}$ such that $|\rho| < t$, in the $(\varphi, \mathcal{A})$-game of Prover and Delayer with starting position ρ there is a linearly described strategy of Delayer that guarantees him to earn at least $t - |\rho|$ white coins while paying at most c black coins. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget. Then any* Res($\oplus$) *refutation of $\varphi \circ g$ has either size at least 2^c or depth at least $\frac{t}{2} \log_{\ell+2}\left(\frac{t}{2c}\right)$.*

To prove Theorem 1.9, we also use the random walk and try to use the same approach we used for regular Res($\oplus$). We take a random full assignment σ and consider a path of length t in the proof graph of $\varphi \circ g$ starting in the empty clause, and at each step, we go to a premise that is

falsified by σ. By Theorem 1.5, with noticeable probability, the random walk reaches a *good* clause C (i.e., $\neg C$ corresponds to an assignment from $\mathcal{A}$). In regular cases, a good clause must contain many linear independent forms. If, for some reason, all good clauses on the distance t from the empty clause indeed have many linearly independent forms, then we can use the same argument we used in the regular case to show the lower bound. So the main case is then on the distance t from the root, there is a good clause C having a small rank of $\neg C$. In this case, we start another random work defined by the random assignment falsifying $\neg C$ of some smaller length t_1. And we continue the same reasoning. Either all good clauses have a large rank of their negations, which implies the lower bound on the proof size, or we can start the next random walk, etc. So we get that either the proof is large or has depth at least $t + t_1 + \dots$.

By applying Theorem 1.9 together with the natural Delayer's strategy on Tseitin formulas, we obtain Theorem 1.4.

Organisation of the paper. In Section 2, we give the basic definitions and facts, including the definitions of closure and lifting. In Section 3, we demonstrate our lifting technique by proving resolution width to Res($\oplus$) rank lifting theorem. In Section 4, we define Prover-Delayer games and prove the random walk theorem for the lifted formula's refutation. In Section 5, we prove the lifting theorem from strategies in advanced (Φ, q)-games to the size of regular Res($\oplus$) refutations. In Section 6, we prove the lifting from resolution depth to regular Res($\oplus$) size. In Section 7, we prove the improved separation between regular Res($\oplus$) and Resolution. In Section 8, we prove size depth tradeoff for Res($\oplus$). In Section 9, we formulate open questions. Dependencies between sections are illustrated in Fig. 2.

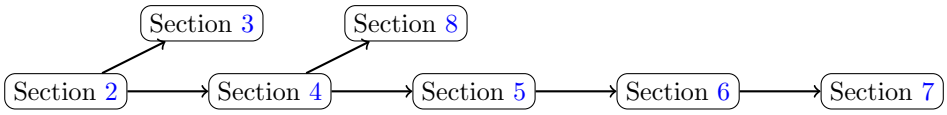

Figure 2: Dependencies between sections.

2 Preliminaries

2.1 Basic notations

For a propositional formula ϕ we denote by $\mathrm{Vars}(\phi)$ the set of all variables mentioned in ϕ. For a set of vectors U from a vector space V we denote by $\langle U \rangle$ the linear span of U.

In this paper, all scalars are from the field $\mathbb{F}_2$. Let X be a set of variables that take values in $\mathbb{F}_2$. A linear form in variables from X is a homogeneous linear polynomial over $\mathbb{F}_2$ in variables from X or, in other words, a polynomial $\sum_i^n x_i a_i$, where $x_i \in X$ is a variable and $a_i \in \mathbb{F}_2$ for all $i \in [n]$. A linear equation is an equality $f = a$, where f is a linear form and $a \in \mathbb{F}_2$.

A *linear clause* is a disjunction of $\mathbb{F}_2$-linear equations: $\bigvee_{i=1}^t (f_i = a_i)$, where f_i are non-zero linear forms, $a_i \in \mathbb{F}_2$, $t \geq 0$ is integer number. Notice that over $\mathbb{F}_2$ a linear clause $\bigvee_{i=1}^t (f_i = a_i)$ may be represented as the negation of a linear system: $\neg \bigwedge_{i=1}^t (f_i = a_i + 1)$.

For a linear clause C we denote by $L(C)$ the set of linear forms that appear in C; i.e. $L\left(\bigvee_{i=1}^t (f_i = a_i)\right) = \{f_1, f_2, \dots, f_t\}$. The same notation we use for linear systems: if Ψ is a $\mathbb{F}_2$-linear

system, $L(\Psi)$ denotes the set of all linear forms from Ψ.

2.2 Resolution over Parities

Let φ be an unsatisfiable CNF formula. A refutation of φ in the proof system Res($\oplus$) [23] is a sequence of linear clauses $C_1, C_2, \ldots, C_s$ such that C_s is the empty clause (i.e., identically false) and for every $i \in [s]$ the clause C_i is either a clause of φ or is obtained from previous clauses by one of the following inference rules:

- *Resolution rule*: From the linear clauses $C \vee (f = a)$ and $D \vee (f = a + 1)$, we can derive the linear clause $C \vee D$.
- *Weakening rule*: From a linear clause C, we can derive any linear clause D in the variables of φ that semantically follows from C, meaning that every assignment satisfying C also satisfies D.

A resolution refutation of a formula φ is a special case of a Res($\oplus$) refutation, where all linear clauses are ordinary clauses.

Any Res($\oplus$) refutation Π of a CNF formula φ can be represented as a directed acyclic graph G_Π with one source. Each node of G_Π is labeled with a linear clause, the source is labeled with the empty clause, sinks are labeled with clauses of ϕ and every node except sinks has one or two outgoing edges such that (1) if a node labeled with C_1 has two outgoing edges to nodes labeled with C_2 and C_3, then C_1 is the result of the resolution rule applied to C_2 and C_3 and (2) if a node labeled with C_1 has only one outgoing edge to a node labeled with C_2, then C_1 is the result of the weakening rule applied to C_2.

We will use another graph $\tilde{G}_\Pi$ obtained from G_Π by contractions of all edges corresponding to weakening rules. For every node u of $\tilde{G}_\Pi$:

- Let u be the result of merging the nodes $v_1, v_2, \ldots, v_k$ $(k > 1)$ forming a path in G_π such that each of the edges $(v_1, v_2), \ldots, (v_{k-1}, v_k)$ of the path corresponds to an application of the weakening rule. Assume that the nodes $v_1, v_2, \ldots, v_k$ are labeled with $C_1, C_2, \ldots, C_k$, respectively;
- We label u with C_k, the strongest of the clauses.

We call the resulting graph $\tilde{G}_\Pi$ *the refutation graph*. It has the following properties:

- $\tilde{G}_\Pi$ is a directed acyclic graph with one source, and each of its sinks is labeled with a clause of φ;
- every node of $\tilde{G}_\Pi$ except sinks has two outgoing edges, and if a node labeled with C_1 has two outgoing edges to nodes labeled with C_2 and C_3, then C_1 is the result of the resolution rule applied to a weakening of C_2 and a weakening of C_3.

By the *size* of a Res($\oplus$) refutation Π, we mean the number of vertices in its refutation graph $\tilde{G}_\Pi$. The *depth* of a Res($\oplus$) refutation Π is the length of the longest path in its refutation graph $\tilde{G}_\Pi$.

2.3 Res($\oplus$) Refutations as Linear Branching Programs

Let X be a set of variables. A *linear branching program* is a directed acyclic graph with one source; every node except sinks has two outgoing edges; for every non-sink node v there is a linear form f_v in variables from X that is called a *query* at the node v; one edge leaving v is labeled $f_v = 0$ and the other edge is labeled $f_v = 1$. Each sink of the graph is labeled with an element from a set A (the set of answers). Every linear branching program computes a function from $\{0,1\}^X \to A$: a full assignment of variables from X determines the unique path from the source to a sink such that this assignment satisfies all equations labeling the path's edges. The sink label is the result of the function.

For every unsatisfiable CNF formula φ we define a relation Search(φ) that consists of all pairs of (σ, C), where σ is an assignment of the variables of φ and C is a clause of φ falsified by σ. We may think of Search(φ) as a search problem where, given an assignment σ, we have to find C such that $(\sigma, C) \in \text{Search}(\phi)$.

Consider a Res($\oplus$) refutation graph G_Π of a CNF formula φ. We now show that the graph G_Π can be relabeled to turn into a linear branching program with the set of answers equal to the set of clauses of φ. Sinks of G_Π are already labeled with clauses of φ. For every non-sink node v of G_π, there is a linear form f_v that is used in the resolution rule at the node v; f_v will be a query at the node v of the linear branching program. Consider an arbitrary node v_1 of G_Π with outgoing edges to nodes v_2 and v_3 and let us define labels of the edges (v_1, v_2) and (v_1, v_3). Let v_1, v_2 and v_3 be labeled with linear clauses C_1, C_2 and C_3, respectively. Let C_1 be the result of the resolution rule applied to $D_2 \vee (f_{v_1} = a)$ and $D_3 \vee (f_{v_1} = a + 1)$, where $D_2 \vee (f_{v_1} = a)$ is a weakening of C_2 and $D_3 \vee (f_{v_1} = a + 1)$ is a weakening of C_3. We label the edge (v_1, v_2) with the linear equation $f_{v_1} = a + 1$ and the edge (v_1, v_3) with $f_{v_1} = a$.

Lemma 2.1 ([13]). *Consider a* Res($\oplus$) *refutation graph with its edges labeled as in the linear branching program associated with it. Let u and v be two nodes labeled with linear clauses C_u and C_v such that a path p connects u to v. Let Φ_p be the conjunction of the equations labeling the edges of p. Then $\Phi_p \wedge \neg C_u$ implies $\neg C_v$. In particular, for any path from the source of a* Res($\oplus$) *refutation graph to a node v labeled with C_v, the system of linear equations written on the edges of this path implies $\neg C_v$.*

Lemma 2.1 implies that every Res($\oplus$) refutation graph of a formula φ may also be considered a linear branching program solving the search problem Search(φ).

For a node v of a linear branching program, we denote by Post(v) the linear span of all linear forms f such that f is a query at a node on a path from v to a sink.

A Res($\oplus$) refutation is called *bottom-regular*, or just *regular*, if for every edge (v, w) in the associated linear branching program $f_v \notin \text{Post}(w)$, where f_v is the query at v.

Lemma 2.2 ([13]). *Suppose that ϕ is an unsatisfiable CNF formula in n variables, and Π is a regular* Res($\oplus$) *refutation of ϕ. Let G_Π be the refutation graph associated with Π. Then, for every node v in G_Π such that there is a path from the source to v of length d, the dimension of* Post(v) *is at most $n - d$.*

For a node v of a linear branching program, we denote by Pre(v) the linear span of all linear forms f such that f is a query at a node $u \neq v$ on a path from the source to v. A Res($\oplus$) refutation is called *top-regular* if for all non-sink nodes v of the linear branching program associated with the refutation, $f_v \notin \text{Pre}(v)$, where f_v is a query at a node v.

A Res($\oplus$) refutation is called *weakly regular* if, for every path from the source to the sink of the linear branching program associated with the refutation, all queries are linearly independent.

Proposition 2.3. *1. Top-regular and bottom-regular resolution refutations are also weakly regular. 2. The depth of any weakly regular resolution refutation of φ is at most* $|\mathrm{Vars}(\varphi)|$.

Proof. 1. Consider a linear branching program associated with a bottom-regular or top-regular refutation and a path $v_1, v_2, \ldots, v_s$ from the source to a sink in it. Let $f_1, f_2, \ldots, f_{s_1}$ are queries in this path. In bottom-regular case for every $i \in [s-1]$, $f_i \notin \mathrm{Post}(v_{i+1})$, hence $f_i \notin \langle f_j \mid s \geq j > i \rangle$. In top-regular case for every $i \in [s-1]$, $f_i \notin \mathrm{Pre}(v_i)$, hence $f_i \notin \langle f_j \mid 1 \leq j < i \rangle$. Thus, in both cases, all $f_1, \ldots, f_s$ are linearly independent. Thus, the refutation is also weakly regular.

2. Consider a linear branching program associated with a weakly regular refutation of φ and a path $v_1, v_2, \ldots, v_s$ from the source to a sink in it. Let $f_1, f_2, \ldots, f_{s_1}$ are queries in this path. The length of the path $s - 1 = \dim\langle f_1, f_2, \ldots, f_{s-1} \rangle \leq |\mathrm{Vars}(\varphi)|$. □

2.4 Safe and Dangerous Sets of Linear Forms

We consider the set of propositional variables $X = \{x_{i,j} \mid i \in [m], j \in [\ell]\}$. The variables from X are divided into m blocks by the value of the first index. The variables $x_{i,1}, x_{i,2}, \ldots, x_{i,\ell}$ form the ith block, for $i \in [m]$.

Consider sets of linear forms using variables from X over the field $\mathbb{F}_2$. The *support* of a linear form $f = x_{i_1,j_1} + x_{i_2,j_2} + \cdots + x_{i_k,j_k}$ is the set $\{i_1, i_2, \ldots, i_k\}$ of blocks of variables that appear in f with non-zero coefficients. We denote the support by $\mathrm{supp}(f)$. The support of a set of linear forms F is the union of the supports of all linear forms in this set. We denote it by $\mathrm{supp}(F)$. We say that a linearly independent set of linear forms F is *dangerous* if $|F| > |\mathrm{supp}(F)|$. We say that a set of linear forms F is *safe* if $\langle F \rangle$ does not contain a dangerous set. If F is linearly dependent but $\langle F \rangle$ contains a dangerous set, instead of saying that F is dangerous, we say it is not safe.

Every linear form corresponds to a vector of its coefficients indexed by the variables from the set X. Given a list of linear forms $f_1, f_2, \ldots, f_k$, one may consider their coefficient matrix of size $k \times |X|$ in which the i-th row coincides with the coefficient vector of f_i.

Theorem 2.4 ([13])**.** *Let $f_1, f_2, \ldots, f_k$ be linearly independent linear forms and let M be their coefficient matrix. Then, the following conditions are equivalent.*

(1) The set of linear forms $f_1, f_2, \ldots, f_k$ is safe.

(2) One can choose k blocks and one variable from each of these blocks such that the columns of M corresponding to the k chosen variables are linearly independent.

2.5 Closure

Let $S \subseteq [m]$ be a set of blocks; for a linear form f we denote by $f[\backslash S]$ a linear form obtained from f by substituting 0 for all variables with support in S. For a set of linear forms F we will use the notation $F[\backslash S] = \{f[\backslash S] \mid f \in F\}$.

A *closure* of a set of linear forms F is any inclusion-wise minimal set $S \subseteq [m]$ such that $F[\backslash S]$ is safe.

Lemma 2.5 (Uniqueness [13])**.** *For any F, its closure is unique.*

We denote the closure of F by $\mathrm{Cl}(F)$.

Lemma 2.6 (Monotonicity [13]). *If $F_1 \subseteq F_2$, then $\mathrm{Cl}(F_1) \subseteq \mathrm{Cl}(F_2)$.*

Lemma 2.7 (Span invariance [13]). $\mathrm{Cl}(F) = \mathrm{Cl}(\langle F \rangle)$.

Lemma 2.8 (Size bound [13]). $|\mathrm{Cl}(F)| + \dim\langle F[\setminus \mathrm{Cl}(F)]\rangle \leq \dim\langle F\rangle$, *and hence* $|\mathrm{Cl}(F)| \leq \dim\langle F\rangle$.

2.6 Tseitin Formulas

Let $G(V, E)$ be a graph. Let $c\colon V \to \{0,1\}$ be a *charge function*. A *Tseitin formula* $\mathrm{T}(G, c)$ depends on the propositional variables x_e for $e \in E$. For each vertex $v \in V$, we define the parity condition of v as $P(v) := \left(\sum_{e \ni v} x_e \equiv c(v) \bmod 2\right)$, where $e \ni v$ means that an edge e is incident to the vertex v. The Tseitin formula $\mathrm{T}(G, c)$ is the conjunction of vertices' parity conditions: $\bigwedge_{v \in V} P(v)$. Tseitin formulas are represented in CNF as follows: we represent $P(v)$ in CNF in a canonical way for all $v \in V$.

Assume that G consists of connected components $H_1, H_2, \ldots, H_t$. Then the Tseitin formula $\mathrm{T}(G, c)$ is equivalent to the conjunction $\bigwedge_{i=1}^{t} \mathrm{T}(H_i, c)$. In the last formula, we abuse the notation since c is defined not only on the vertices of H_i; thus, we implicitly use the corresponding restriction on the set of vertices.

Lemma 2.9 (Folklore, see e.g. [29]). *A Tseitin formula $\mathrm{T}(G, c)$ is satisfiable if and only if for every connected component $C(U, E_U)$ of the graph G, the condition $\sum_{u \in U} c(u) \equiv 0 \bmod 2$ holds.*

Corollary 2.10. *Let $G(V, E)$ be connected graph and $c : V \to \{0,1\}$. Then, for every $u \in V$, the conjunction of parity conditions for all vertices except u, i.e., $\bigwedge_{v \in V, v \neq u} P(v)$ is satisfiable.*

Proof. Let $c' : V \to \{0,1\}$ differ from from c only in u. By Lemma 2.9 the one formula from $\mathrm{T}(G, c)$ and $\mathrm{T}(G, c')$ is satisfiable. The formula $\bigwedge_{v \in V, v \neq u} P(v)$ is a subformula of both of them; hence it is also satisfiable. □

Lemma 2.11 (Folklore). *The result of the substitution $x_e := b$ to $\mathrm{T}(G, c)$ where $b \in \{0,1\}$ is a Tseitin formula $\mathrm{T}(G', c')$ where $G' = G - e$ and c' differs from c on the endpoints of the edge e by b and equals c for every other vertex.*

2.7 Spectral Expanders

Let $G(V, E)$ be an undirected graph without loops but possibly with multiple edges. G is a spectral (n, d, α)-expander if G is d-regular, $|V| = n$, and the absolute value of the second largest eigenvalue of the adjacency matrix of G is not greater than αd.

It is well known that for all $1 > \alpha > 0$ and all large enough constants d there exist natural number n_0 and a family $\{G_n\}_{n=n_0}^{\infty}$ of (n, d, α)-expanders. There are explicit constructions such that G_n can be constructed in $\mathrm{poly}(n)$ time [25]. Also, it is known that a random d-regular graph is an expander with high probability.

Let us denote by $E(A, B)$ a multiset of edges with one end in A and another in B. Note that when both ends of an edge are simultaneously in A and in B, we count this edge twice.

Lemma 2.12 (Cheeger inequality [10]). *Let $G(V, E)$ be an (n, d, α)-expander. Then for all $A \subseteq V$ such that $|A| \leq \frac{n}{2}$ the following inequality holds: $|E(A, V \setminus A)| \geq \frac{1-\alpha}{2} d|A|$.*

Corollary 2.13. *Every (n, d, α)-expander with $0 < \alpha < 1$ is connected.*

Proof. If G is not connected, then we will get a contradiction with Lemma 2.12 if we choose A to be the smallest connected component. □

Lemma 2.14 (Expander mixing lemma [4])**.** *Let $G(V, E)$ be (n, d, α)-expander, $A, B \subseteq V$. Then $\left||E(A, B)| - \frac{d|A||B|}{n}\right| \leq \alpha d\sqrt{|A||B|}$.*

Lemma 2.15 (Lemma 11 from [16])**.** *Every graph that can be obtained by deleting $l \leq \frac{n}{4}$ edges from an algebraic (n, d, α)-expander G contains at most $\frac{2l}{d(1-\alpha)} + 1$ connected components.*

2.8 Lifting of Formulas via Gadget

For every CNF formula φ with variables $Y = \{y_1, y_2, \ldots, y_m\}$ and every Boolean function $g : \{0,1\}^\ell \to \{0,1\}$ we define a CNF formula $\varphi \circ g$ with variables $X = \{x_{i,j} \mid i \in [m], j \in [\ell]\}$ representing $\varphi(g(x_{1,1}, x_{1,2}, \ldots, x_{1,\ell}), g(x_{2,1}, x_{2,2}, \ldots, x_{2,\ell}), \ldots, g(x_{m,1}, x_{m,2}, \ldots, x_{m,\ell}))$ (i.e. we substitute to every variable of φ the function g applied to ℓ fresh variables). Let $\varphi = \bigwedge_{i \in I} C_i$, where C_i is a clause for all $i \in I$. For every $i \in [m]$ we denote by $y_i \circ g$ the cannonical CNF formula representing $g(x_{i,1}, x_{i,2}, \ldots, x_{i,\ell})$ which has ℓ variables in every clause and by $(\neg y_i) \circ g$ the cannonical CNF formula representing $\neg g(x_{i,1}, x_{i,2}, \ldots, x_{i,\ell})$ which has ℓ variables in every clause. Let $C_i = l_{i,1} \vee l_{i,2} \vee \cdots \vee l_{i,n_i}$, where $l_{i,j}$ is a literal. Then we denote by $C_i \circ g$ a CNF formula that represents $l_{i,1} \circ g \vee l_{i,2} \circ g \vee \cdots \vee l_{i,n_i} \circ g$ as follows: $C_i \circ g$ consists of all clauses of the form $D_1 \vee D_2 \vee \cdots \vee D_{n_i}$, where D_j is a clause of $l_{i,j} \circ g$ for all $j \in [n_i]$. And $\varphi \circ g := \bigwedge_{i \in I} C_i \circ g$.

Lemma 2.16. *If a clause C contains variables $\{y_i\}_{i \in I}$, then every clause of $C \circ g$ contains variables $\{x_{i,j} \mid i \in I, j \in [\ell]\}$.*

Proof. The definition straightforwardly implies the lemma. □

We refer to $\varphi \circ g$ as a formula φ *lifted with a gadget* g. We refer to the set $Y = \{y_1, y_2, \ldots, y_m\}$ as a set of *unlifted* variables and to the set $X = \{x_{i,j} \mid i \in [m], j \in [\ell]\}$ as a set of *lifted* variables.

Sometimes, we will identify subsets of $[m]$ with corresponding subsets of Y. It is especially convenient to use such correspondence for the notions of support and closure. So, we will assume that the support and the closure of the set of linear forms over lifted variables is the set of unlifted variables.

A partial assignment ρ to the set of variables X is called block-respectful if, for every i, ρ either assigns values to all variables with support i or does not assign values to any of them.

Suppose that ρ is a block-respectful partial assignment. Then we define by $\hat{\rho}$ the partial assignment on the set of variables Y such that $\hat{\rho}(y_i) = g(\rho(x_{i,1}), \rho(x_{i,2}), \ldots, \rho(x_{i,\ell}))$ (here we assume that if the right-hand side is undefined, then the left-hand side is also undefined).

Let $k < \ell$. A gadget (i.e. Boolean function) $g : \{0,1\}^\ell \to \{0,1\}$ is called *k-stifling* [9] if for every $A \subset [\ell]$ of size k for every $c \in \{0,1\}$ there exists $a \in \{0,1\}^\ell$ such that for every $b \in \{0,1\}^\ell$ if a and b agree on set of indices $[\ell] \setminus A$, then $g(b) = c$.

Lemma 2.17. *Let Ψ be a satisfiable linear system in the lifted variables X and $L(\Psi)$ be safe. Let $g : \{0,1\}^\ell \to \{0,1\}$ be an 1-stifling gadget. Then for any full assignment σ to the unlifted variables Y there exists a full assignment τ to the lifted variables X such that τ satisfies Ψ and $\hat{\tau} = \sigma$.*

Proof. W.l.o.g. assume that all equations of Ψ are linearly independent.

By Theorem 2.4, the span of columns of the coefficient matrix of Ψ contains a basis that contains at most one column for every block. Let $Z \subseteq X$ be the set of variables corresponding to this basis. Using a 1-stifling property of g we can construct the full assignment τ_0 to variables $X \setminus Z$ such that for every assignment ρ extending τ_0 to the set of variables X, for every $i \in [m]$, $g(\rho(x_{i,1}), \rho(x_{i,2}), \dots, \rho(x_{i,\ell})) = \sigma(y_i)$. Since Z corresponds to the basis of the span of the columns of the linear system Ψ, τ_0 can be extended to an assignment τ that satisfies Ψ. By the construction, $\hat{\tau} = \sigma$. □

The following lemma appeared in [7]; we provide its proof since it is a simple corollary of Lemma 2.17. Informally, the next lemma states that any linear system over the lifted variables restricts unlifted variables only from the closure.

Lemma 2.18 (Lemma 17 from [7])**.** *Let Ψ be a satisfiable linear system in the lifted variables X. Let $g : \{0,1\}^\ell \to \{0,1\}$ be an 1-stifling gadget. Suppose*

- *σ is a full assignment to lifted variables X satisfying Ψ.*
- *π is a full assignment to unlifted variables Y such that $\pi|_{\mathrm{Cl}(L(\Psi))} = \hat{\sigma}|_{\mathrm{Cl}(L(\Psi))}$.*

Then there exists a full assignment τ to the lifted variables X such that τ satisfies Ψ and $\hat{\tau} = \pi$.

Proof. Let T be the set of all lifted variables with support in $\mathrm{Cl}(L(\Psi))$. Let σ_0 be the restriction of σ to T. The linear system $(\Psi)|_{\sigma_0}$ is satisfiable, and its set of linear forms is safe by the definition of closure. By Lemma 2.17, there exists an assignment γ to the lifted variables $\mathrm{Vars}(\Psi) \setminus T$ that satisfies $(\Psi)|_{\sigma_0}$ and such that $\widehat{(\sigma_0 \cup \gamma)} = \pi$. Thus, we can take $\tau = \sigma_0 \cup \gamma$. □

Lemma 2.19. *Let Ψ be a satisfiable linear system in the lifted variables X. Let $g : \{0,1\}^\ell \to \{0,1\}$ be an 1-stifling gadget. Suppose there exists a full assignment σ to lifted variables X satisfying Ψ such that $\hat{\sigma}|_{\mathrm{Cl}(L(\Psi))}$ does not falsify any clause of φ. Then, Ψ does not contradict any clause of $\varphi \circ g$.*

Proof. Consider a clause C' from $\varphi \circ g$ and a clause C from φ such that C' is a clause from $C \circ g$.

Since $\hat{\sigma}|_{\mathrm{Cl}(L(\Psi))}$ does not falsify any clause of φ, there exists a full assignment of unlifted variables π that extends $\hat{\sigma}|_{\mathrm{Cl}(L(\Psi))}$ and satisfies C. By Lemma 2.18, there exists an assignment τ of the lifted variables that satisfies Ψ and such that $\hat{\tau} = \pi$. Hence, τ satisfies $C \circ g$ and, thus, τ satisfies C'. □

3 Lifting resolution width to Res(⊕) rank

The *width* of a resolution refutation is the maximal number of literals in any clause of the refutation.

Similarly, we can define the rank of Res(⊕) refutation. The *rank* of a Res(⊕) refutation is the maximal rank of the negation of any linear clause in the refutation.

The goal of this section is to relate these two measures by a lifting theorem:

Theorem 3.1. *Let $g\colon \{0,1\}^\ell \to \{0,1\}$ be a 1-stifling gadget. Consider a CNF formula φ such that $\varphi \circ g$ has a* Res(⊕) *refutation of rank W. Then φ should have a resolution refutation of width at most W.*

To prove this theorem, we will need to consider Spoiler-Duplicator games.

Spoiler-Duplicator game for resolution. Firstly defined by Atserias and Dalmau [5], the k-pebble game on an unsatisfiable CNF formula φ proceeds as follows: starting with the empty assignment, on every turn Spoiler has two options:

- If the size of the current assignment is less than k, then Spoiler can ask Duplicator about the value of some variable x from φ. Then Duplicator chooses the value of x.
- Spoiler can erase one of the variables from the domain of the assignment.

Spoiler wins if the current assignment contradicts one of the clauses of φ. Duplicator wins if he can answer Spoiler's responses such that Spoiler does not win.

One can formalize this game by defining a Duplicator's winning strategy as follows:

Let φ be an unsatisfiable CNF formula. We say that the Duplicator wins the k-pebble game on φ if there is a non-empty family $\mathcal{H}$ of partial truth assignments that do not falsify any clause from φ such that:

- If $f \in \mathcal{H}$, then $|f| \leq k$.
- If $f \in \mathcal{H}$ and $g \subseteq f$, then $g \in \mathcal{H}$.
- If $f \in \mathcal{H}$, $|f| < k$ and x is a variable, then there is value $a \in \{0,1\}$ such that $f \cup \{x := a\} \in \mathcal{H}$.

The following equivalence establishes the connection between the resolution width and Spoiler-Duplicator games:

Lemma 3.2 ([5]). *Let φ be a k-CNF formula, and W be an integer number. Then φ does not have a resolution refutation of width W if and only if a Duplicator's winning strategy exists in $(W+1)$-pebble game on φ.*

Spoiler-Duplicator game for Res($\oplus$)**.** Similarly to the games of Atserias and Dalmau, one can define Spoiler-Duplicator k-pebble Res($\oplus$)-games [18]. This game on an unsatisfiable CNF formula φ proceeds as follows: starting with the empty system of $\mathbb{F}_2$-linear equations, on every turn, Spoiler has two options:

- If the rank of the current system is less than k, then Spoiler can ask Duplicator about the value of some *linear form* ℓ over the variables from φ. Then Duplicator can choose the value $a \in \{0,1\}$ of ℓ and add $\ell = a$ to the current linear system.
- Spoiler can change the current linear system to any other linear system, which is implied by the current system.

Spoiler wins if the current system contradicts one of the clauses of φ. Duplicator wins if he can answer Spoiler's responses such that Spoiler does not win.

Similarly to [18], we formally define a Duplicator's strategy in Res($\oplus$)-games: We say that the Duplicator wins the Res($\oplus$) existential k-pebble game on φ if there is a non-empty family $\mathcal{H}$ of linear systems over $\mathbb{F}_2$ such that:

- For every $F \in \mathcal{H}$ and every clause C in φ, there exists a solution of F that satisfies C.
- If $F \in \mathcal{H}$, then $\mathrm{rk}(F) \leq k$.

- If $F \in \mathcal{H}$ and F semantically implies G, then $G \in \mathcal{H}$.
- If $F \in \mathcal{H}$ and $\mathrm{rk}(F) \leq k-1$ and f is a linear form, then there is $a \in \mathbb{F}_2$ such that $F \wedge \{f = a\} \in \mathcal{H}$.

Similarly to Lemma 3.2, one can prove the following lemma:

Lemma 3.3 (cf. [18])**.** *If for an unsatisfiable CNF formula φ Duplicator has a winning strategy $\mathcal{H}$ in a $(W+1)$-pebble* Res($\oplus$)*-game, then φ does not have a* Res($\oplus$) *refutation of rank at most W.*

Proof. Suppose there is a Res($\oplus$) refutation Π for φ of rank at most W and a Duplicator's winning strategy $\mathcal{H}$ in the $(W+1)$-pebble Res($\oplus$)-game. Consider a linear branching program associated with Π. Consider the following strategy of Spoiler: it starts at the source labeled with the empty clause, and the current linear system is also empty. Every his move, Spoiler asks for a value of the linear form corresponding to the current node (this move is legal since the rank of the current linear system is at most W) and moves to the node corresponding to the answer of Duplicator and then changes the current linear system to the negation of the clause in the new node (this move is legal by Lemma 2.1). Properties of the Duplicator's strategy imply that this process will never stop since the current linear system does not contradict clauses of φ. This is a contradiction since there are no infinite paths in the directed acyclic graph. □

Now, using those two lemmas, we can prove the main result of this section.

Proof of Theorem 3.1. Suppose there is no resolution refutation of φ of width W. Consider a $(W+1)$-winning strategy $\mathcal{H}$ for the resolution Spoiler-Duplicator games for φ that exists by Lemma 3.2. Consider the following family of linear systems $\tilde{\mathcal{H}}$: it will consist of all $\mathbb{F}_2$-linear systems F over the variables from $\varphi \circ g$ for which the following holds:

- $\mathrm{rk}(F) \leq W+1$.
- There exist $h \in \mathcal{H}$ and a solution σ of F such that $\hat{\sigma}$ coincides with h on $\mathrm{Cl}(L(F))$.

We show that $\tilde{\mathcal{H}}$ satisfies all the properties of $(W+1)$-winning Res($\oplus$)-strategy:

- By definition, for every $F \in \tilde{\mathcal{H}}$, $\mathrm{rk}(F) \leq W+1$.
- By Lemma 2.19, for every $F \in \tilde{\mathcal{H}}$ and every clause C' from $\varphi \circ g$, there exists a solution of F that satisfies C'.
- Let us show that if G is a linear system satisfying and for some $F \in \tilde{\mathcal{H}}$, F semantically implies G, then $G \in \tilde{\mathcal{H}}$. Indeed, since F semantically implies G, $L(G) \subseteq \langle L(F) \rangle$. Then by Lemmas 2.6 and 2.7, $\mathrm{Cl}(L(G)) \subseteq \mathrm{Cl}(L(F))$. Clear that $\mathrm{rk}(G) \leq \mathrm{rk}(F) \leq W+1$. Since, $F \in \tilde{\mathcal{H}}$ there exist $h \in \mathcal{H}$ and there is a solution σ of F such that $\hat{\sigma}$ coincides with h on $\mathrm{Cl}(L(F))$. Notice that σ is also a solution of G and $\hat{\sigma}$ coincides with h on $\mathrm{Cl}(L(G))$. Hence G is in $\tilde{\mathcal{H}}$.
- Finally, we need to show that for any $F \in \tilde{\mathcal{H}}$ with $\mathrm{rk}(F) < W+1$ and for every linear form f, there exists a constant $a \in \mathbb{F}_2$ such that $F \wedge \{f = a\} \in \tilde{\mathcal{H}}$.

  There exist $h \in \mathcal{H}$ and a solution σ of F such that $\hat{\sigma}$ coincides with h on $\mathrm{Cl}(L(F))$. W.l.o.g., assume that the domain of h is precisely $\mathrm{Cl}(L(F))$. By Lemma 2.8, $|(\mathrm{Cl}(L(F) \cup \{f\}))| \leq$

$\mathrm{rk}(F) + 1 \leq W + 1$. By the properties of $\mathcal{H}$ there is $g \in \mathcal{H}$ such that $h \subseteq g$ and g is defined on $\mathrm{Cl}(L(F) \cup \{f\})$; indeed, we can extend h for all variables from $\mathrm{Cl}(L(F) \cup \{f\})) \setminus \mathrm{Cl}(L(F))$ one by one. Using Lemma 2.18, we can find a solution τ of F such that $\hat{\tau}$ coincides with g on $\mathrm{Cl}(L(F) \cup \{f\}))$. Let a be a value of linear form f on the solution τ. Then τ clearly satisfies $F \wedge \{f = a\}$. On the other hand, $\hat{\tau}$ coincides with g on $\mathrm{Cl}(L(F) \cup \{f\}))$. Thus $F \wedge \{f = a\} \in \tilde{\mathcal{H}}$.

Since $\tilde{\mathcal{H}}$ is $(W+1)$-winning $\mathrm{Res}(\oplus)$-strategy, by Lemma 3.3, it is impossible to construct a $\mathrm{Res}(\oplus)$ refutation for $\varphi \circ g$ of rank W. □

4 Prover-Delayer Games and Random Walk

4.1 Prover-Delayer Games

Let φ be a CNF formula. Let $\mathcal{A}$ be a set consisting of partial assignments for variables of φ. We define a $(\varphi, \mathcal{A})$-*game of Prover and Delayer with starting position* $\rho_0 \in \mathcal{A}$. In this game, there are two players: Prover and Delayer. The players save the current partial assignment ρ that initially equals ρ_0. On every move, Prover chooses a variable $x \in \mathrm{Vars}(\varphi) \setminus \mathrm{Dom}(\rho)$, and Delayer has two options:

- Delayer can earn a *white* coin and reports $*$. Then, Prover chooses a Boolean value a of x.
- Delayer can earn a *white* coin and pay a *black* coin to choose a Boolean value a of x by himself.

After the move the current assignment is updated: $\rho := \rho \cup \{x := a\}$; The game ends when $\rho \notin \mathcal{A}$ or $\mathrm{Dom}(\rho) = \mathrm{Vars}(\varphi)$.

A strategy of Delayer is a function $f\colon H \times \mathrm{Vars}(\varphi) \to \{0, 1, *\}$, where H is the set of all possible sequences of pairs of queries asked by Prover with answers (so it is the sequences of the form $(x_{i_1} := \alpha_1, x_{i_2} := \alpha_1, \ldots, x_{i_k} := \alpha_k)$). The strategy is utilized in a natural way: on every step, given the sequence of previous queries with answers Q and a last queried variable x, Delayer answers $f(Q, x)$. If $f(Q, x) \neq *$, we say that the value of x is *forced*.

A linearly described strategy is a special case of a strategy of Delayer. A linearly described strategy has the following form: given a sequence of queries $(x_{i_1}, x_{i_2}, \ldots, x_{i_k})$ (without the answers for those queries) and a last queried variable x it outputs either $*$ or an $\mathbb{F}_2$-linear function $h(x_{i_1}, x_{i_2}, \ldots, x_{i_k})$.

Delayer utilizes this strategy in the following way. On every step given a sequence of queries with answers $(x_{i_1} := \alpha_1, x_{i_2} := \alpha_1, \ldots, x_{i_k} := \alpha_k)$ and a last queried variable x: if strategy outputs $*$ on $(x_{i_1}, x_{i_2}, \ldots, x_{i_k}, x)$, then Delayer answers $*$; if strategy outputs function $h(x_{i_1}, x_{i_2}, \ldots, x_{i_k})$, then Delayer answers $h(\alpha_1, \alpha_2, \ldots, \alpha_k)$.

We say that a strategy of Delayer *guarantees him to earn t white coins while paying at most c black coins* if, for every behavior of Prover, Delayer, using his strategy, achieves the position with the current assignment $\rho \in \mathcal{A}$ such that by this moment he earns t white coins and pay at most c black coins.

4.2 Natural Strategy for Tseitin Formulas

In this subsection, we give an example of Prover-Delayer games and the strategy of Delayer for Tseitin formulas based on expanders.

Proposition 4.1. *If a graph H is obtained from an (n,d,α)-expander by the deletion of at most $\frac{dn}{4}\cdot\frac{1-\alpha}{2}$ edges, then the largest connected component of H has size more than $n/2$.*

Proof. Proof by contradiction. Let us delete edges from the (n,d,α) expander one by one. Consider the first moment where the largest connected component becomes at most $n/2$. Its size at this moment is greater than $n/4$. By Lemma 2.12, by this moment we have to remove more than $\frac{dn}{4}\cdot\frac{1-\alpha}{2}$ edges. □

Let $\mathrm{T}(G,c)$ be an unsatisfiable Tseitin formula based on a spectral (n,d,α)-expander $G(V,E)$.

Consider a partial assignment ρ for variables of $\mathrm{T}(G,c)$. By Lemma 2.11, $\mathrm{T}(G,c)|_\rho$ is also a Tseitin formula; we denote by G_ρ and c_ρ the graph and the charge function for $\mathrm{T}(G,c)|_\rho$.Note that the graph G_ρ is obtained from G by the deletion of edges corresponding to $\mathrm{Dom}(\rho)$. Let $U^1_\rho, U^2_\rho, \ldots, U^{k_\rho}_\rho \subseteq V$ be the connected components of G_ρ sorted by decreasing of their sizes: $|U^1_\rho| \geq |U^2_\rho| \geq \cdots \geq |U^{k_\rho}_\rho|$. By Proposition 4.1, if $|\rho| \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$, $|U^1_\rho| > n/2$ and U^1_ρ is the unique largest connected component of G_ρ. It is easy to see that the Tseitin formula $\mathrm{T}(G_\rho,c_\rho)$ is the conjunction of independent Tseitin formulas corresponding to the components of G_ρ and the restrictions of c_ρ to them.

Let $\mathcal{T}$ be the set of all partial assignments ρ for variables of $\mathrm{T}(G,c)$ such that $|\rho| \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$ and $\sum_{v\in U^1_\rho} c_\rho(v) = 1$ and for $i \in [k_\rho]\setminus\{1\}$, $\sum_{v\in U^i_\rho} c_\rho(v) = 0$. In other words, $\rho\in\mathcal{T}$ iff $|\rho| \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$ and the Tseitin formula corresponding to the largest connected component of G_ρ is unsatisfiable, but Tseitin formulas corresponding to all other connected components of G_ρ are satisfiable.

Consider the following strategy of Delayer for the game based on $T(G,c)$ and every starting position $\rho_0 \in \mathcal{T}$. Let ρ be the current assignment in the game. The strategy will maintain the following invariant: if $|\rho| \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$, then $\rho\in\mathcal{T}$. The strategy is the following.

- If $|\rho|+1 \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$ and the request corresponds to a non-bridge edge in the graph of G_ρ, Delayer responds $*$.
  - In this case, connected components of G_ρ are not changed, and the substitution of a non-bridge edge does not reflect their satisfiability.
- If $|\rho|+1 \leq \frac{dn}{4}\cdot\frac{1-\alpha}{2}$ and the request corresponds to a bridge e in the graph of G_ρ, Delayer chooses the answer to satisfy the invariant as follows.
  - Assume that the bridge e connects two vertices from U^i_ρ for $i\in[k_\rho]$ and let the removing e from G_ρ split U^i_ρ to the new connected components A and B and $|A|\geq|B|$. According the stratagy Delayer chooses the value $\gamma = \sum_{v\in B} c_\rho(v)$ to x_e. We denote $\rho' := \rho \cup x_e := \sum_{v\in B} c_\rho(v)$. Below, we verify that the invariant is satisfied in two different cases: $i\neq 1$ and $i=1$.
  - Case $i\neq 1$. Then $\sum_{v\in U^i_\rho} c_\rho(v) = 0$, hence $\sum_{v\in B} c_\rho(v) = \sum_{v\in A} c_\rho(v) = \gamma$. Then $\sum_{v\in B} c_{\rho'}(v) = \sum_{v\in B} c_\rho(v) + \gamma = 0$ and $\sum_{v\in A} c_{\rho'}(v) = \sum_{v\in A} c_\rho(v) + \gamma = 0$. In this case, A and B are not the largest connected component of G_ρ, thus, the invariant holds.

– Case $i = 1$. Then $\sum_{v \in U^i_\rho} c_\rho(v) = 1$. Then $\sum_{v \in B} c_\rho(v) = \gamma$, $\sum_{v \in A} c_\rho(v) = 1 + \gamma$. Then $\sum_{v \in B} c_{\rho'}(v) = \sum_{v \in B} c_\rho(v) + \gamma = 0$ and $\sum_{v \in A} c_{\rho'}(v) = \sum_{v \in A} c_\rho(v) + \gamma = 1$. In this case, A is the largest connected component in $G_{\rho'}$.

- If $|\rho| + 1 > \frac{dn}{4} \cdot \frac{1-\alpha}{2}$, Delayer responds $*$.

We will refer to this strategy as the *natural strategy*.

Lemma 4.2. *The natural strategy is linearly described.*

Proof. Forced variables in the natural strategy correspond to bridges. Let ρ be the current assignment, and Prover asks the value of an edge e, which is a bridge of G_ρ connecting two vertices from the connected component U^i_ρ. Let the removing of e from G_ρ split U^i_ρ to the new components A and B and $|A| \geq |B|$. Then, Delayer, according to the natural strategy, chooses the value $\gamma = \sum_{v \in B} c_\rho(v)$ to x_e.

Let I be set of edges corresponding to variables $\mathrm{Dom}(\rho) \setminus \mathrm{Dom}(\rho_0)$ connecting B with $V \setminus B$. Then $\gamma = \sum_{v \in B} c_\rho(v) = \sum_{v \in B} c_{\rho_0}(v) + \sum_{e \in I} \rho(x_e)$. Hence, γ is the result of an affine function applied to the values of previous variables. □

4.3 Random Walk

We say that a set of partial assignments $\mathcal{A}$ is closed under restrictions, if for every $\rho \in \mathcal{A}$ for every $\sigma \subseteq \rho$, $\sigma \in \mathcal{A}$.

Theorem 4.3. *Let φ be an unsatisfiable CNF formula. Let $\mathcal{A}$ be a set of partial assignments for $\mathrm{Vars}(\varphi)$ such that $\mathcal{A}$ is closed under restrictions and for any $\sigma \in \mathcal{A}$, σ does not falsify any clause of φ. Assume that in the $(\varphi, \mathcal{A})$-game, Delayer has a linearly described strategy with start position $\rho_0 \in \mathcal{A}$ that guarantees him to earn w white coins while paying at most c black coins. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget. Consider a $\mathrm{Res}(\oplus)$ refutation of $\varphi \circ g$ and the linear branching program associated with it. Let C_0 be a linear clause from this refutation. Assume that $\mathrm{Cl}(L(C_0)) = \mathrm{Dom}(\rho_0)$ and there is a solution τ of $\neg C_0$ such that $\hat{\tau}$ extends ρ_0. Let Σ be the set of all full assigmnets π such that π satisfies $\neg C_0$ and $\hat{\pi}$ extends ρ_0. Let t be integer number such that $t \leq w - \mathrm{rk}(\neg C_0) + |\rho_0|$. (If C_0 is an empty clause, then ρ_0 is an empty assignment; hence, in this case, Σ is the set of all assignments, and t is restricted to be at most w.) Consider a random full assignment $\sigma \in \Sigma$ and make t steps in the linear branching program from C_0 according to σ (if we reach a sink earlier than in t steps, we stay there); let we stop in a node labeled with a linear clause C. Then $\hat{\sigma}|_{\mathrm{Cl}(L(C))} \in \mathcal{A}$ with probability at least $2^{-c(\ell-1)}$.*

We start with an informal proof plan of Theorem 4.3.

1. Let Ψ_i denote the conjunction of $\neg C_0$ and the linear system corresponding to the first i edges of the path in the linear branching program defined by σ. By the properties of linear branching program $\Psi_t \models \neg C$ and, hence, $L(C) \subseteq \langle L(\Psi_t) \rangle$ and, thus by Lemmas 2.6 and 2.7, $\mathrm{Cl}(L(C)) \subseteq \mathrm{Cl}(L(\Psi_t))$. Since $\mathcal{A}$ is closed under restrictions, $\Pr[\hat{\sigma}|_{\mathrm{Cl}(L(C))} \in \mathcal{A}] \geq \Pr[\hat{\sigma}|_{\mathrm{Cl}(L(\Psi_t))} \in \mathcal{A}]$. Let us denote the latter probability by P^*. We will prove that $P^* \geq 2^{-c(\ell-1)}$.

2. To estimate P^* w.l.o.g. we may assume that the linear branching program is a tree (i.e., parity decision tree). We can convert a linear branching program to a tree in a standard way by repeating the nodes at the same distance from the source.

3. We consider the following metaphor: we assume that elements of Σ are grains of sand. Initially, we put all the sand in a node of the parity decision tree labeled with C_0. Each round, every grain of sand in an interior node moves to a child of the current node corresponding to the equation on the edge going to this child. A grain of sand (i.e., full assignment) τ disappears from the node v (and does not move to children in the tree) if $\hat{\tau}|_{\mathrm{Cl}(L(F_v))}$ is not consistent with Delayer's strategy, where F_v is the conjunction of $\neg C_0$ with the system of equations written on the path from C_0 to v. Notice that if $\hat{\tau}|_{\mathrm{Cl}(L(F_v))}$ is consistent with Delayer's strategy, since $|\mathrm{Cl}(L(F_v))| \leq \mathrm{rk}(C_0) + t \leq |\rho_0| + w$, by the properties of the strategy, $\hat{\tau}|_{\mathrm{Cl}(L(F_v))} \in \mathcal{A}$. It is easy to see that P^* is at least the fraction of sand still in the tree in t steps.

4. The following lemma allows us to estimate the fraction of sand still in the tree in t steps.

Lemma 4.4. *Consider a binary tree with root r and a set of leaves L. We associate every node v except leaves with a number $p_v \neq 0$. For every node v of the tree, there is a number n_v such that if u and w are children of v, then $n_v p_v = (n_u + n_w)$. Let for every leaf l the unique path from the root to l be denoted $\pi_l = (s_1 = r, s_2, \ldots, s_t = l)$; let us denote $p(\pi_l) = \prod_{i=1}^{t-1} p_{s_i}$. Then $n_r = \sum_{l \in L} n_l \frac{1}{p(\pi_l)}$.*

Proof. Induction on the number of leaves in the tree. □

In Lemma 4.4, n_v is the amount of sand that was in the node v. And p_v is the fraction of sand that does not disappear when the sand moves from v to its children.

5. $P^* \geq \frac{1}{n_r} \sum_{l \in L} n_l$, where L consists of the nodes on distance t and the leaves on distance at most t. If l is a leaf of the parity decision tree, then F_l contradicts a clause of $\varphi \circ g$. Hence, there is no sand in the leaves, and we can assume that L does not contain leaves. To estimate the probability P^*, it is sufficient to bound from below $p(\pi_l)$ for some node $l \in L$. If we denote by $s_1 = r, s_2, \ldots, s_t = l$ the path from the root to l, then $p(\pi_l) = \prod_{i=1}^{t-1} p_{s_i}$. The following lemma estimates p_{s_i} regarding Delayer's strategy.

Lemma 4.5. *Let h be a linear description of Delayer's strategy in the $(\varphi, \mathcal{A})$-game with starting position ρ_0. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget. Let F be a system of linear equations in the lifted variables (i.e., variables of $\varphi \circ g$). Let f be a linear form. Consider some order θ on unlifted variables such that variables from $\mathrm{Cl}(L(F))$ preceded variables from $\mathrm{Cl}(L(F) \cup \{f\}) \setminus \mathrm{Cl}(L(F))$ that preceded all other variables. Let T be the set of solutions τ of F such that $\hat{\tau}|_{\mathrm{Cl}(L(F))}$ is consistent with the strategy h if variables appear according θ. Let T' be the set of solutions τ of F such that $\hat{\tau}|_{\mathrm{Cl}(L(F) \cup \{f\})}$ is consistent with the strategy h if variables appear according θ. Then $|T'| \geq |T| 2^{-(\ell-1)n}$, where n is the number of variables from $\mathrm{Cl}(L(F) \cup \{f\}) \setminus \mathrm{Cl}(L(F))$ such that Delayer recognizes them as forced according the strategy h if variables appear in the order θ.*

We will prove Lemma 4.5 in Subsection 4.4.

Lemma 4.5 and properties of Delayer's strategy imply that $p(\pi_l) \geq 2^{(\ell-1)c}$ for every $l \in L$. Thus by Lemma 4.4,

$$P^* \geq \frac{\sum_{l \in L} n_l}{n_r} = \frac{\sum_{l \in L} n_l}{\sum_{l \in L} \frac{n_l}{p(\pi_l)}} \geq 2^{-(\ell-1)c}.$$

Proof of Theorem 4.3. We denote by Ψ_i the conjunction of $\neg C_0$ and the linear system corresponding to the first i edges of the path in the linear branching program defined by σ. By Lemma 2.1, the linear system Ψ_t semantically implies $\neg C$, then $L(C) \in \langle L(\Psi_t) \rangle$, then by Lemmas 2.6 and 2.7, $\mathrm{Cl}(L(C)) \subseteq \mathrm{Cl}(L(\Psi_t))$. Since $\mathcal{A}$ is closed under restrictions, it is sufficient to prove that with probability at least $2^{-c(\ell-1)}$, $\hat{\sigma}|_{\mathrm{Cl}(L(\Psi_t))} \in \mathcal{A}$.

We convert the linear branching program to a parity decision tree in the standard way by making several copies of nodes and edges with the same labels. Since the end of the path corresponding σ has the same label in the tree and the linear branching program, we can continue reasoning assuming we walk in the parity decision tree. Let H be the subtree of the parity decision tree rooted in C_0 that contains vertices on the distance at most t from C_0; let r be the root of H.

Let v be a vertex of H. We denote by F_v conjunction of $\neg C_0$ and the system linear equations written on the path from the root to v.

For every v of H we construct an order θ_v of the set of variables $\mathrm{Cl}(L(F_v))$. We construct them by induction from the root to the leaves. θ_r is some order on $\mathrm{Cl}(L(C_0))$. If u and w are children of v, then $\theta_u = \theta_w$ and equals to an order that extends θ_v such that all elements of $\mathrm{Cl}(L(F_v))$ preceded to all elements of $\mathrm{Cl}(L(F_u)) \setminus \mathrm{Cl}(L(F_v))$.

For every vertex v of the tree, we define a set T_v consisting of the set of full assignments τ satisfying F_v such that $\hat{\tau}|_{\mathrm{Cl}(L(F_v))}$ is consistent with Delayer's strategy when variables appear in the order θ_v.

For the root, $T_r = \Sigma$. By Lemma 2.8, for every vertex v of H, $|\mathrm{Cl}(L(F_v))| \leq \dim\langle L(F_v)\rangle \leq t + \mathrm{rk}(\neg C_0) \leq w + |\rho_0|$, hence for every $\sigma \in T_v$, $\hat{\sigma}|_{\mathrm{Cl}(L(F_v))} \in \mathcal{A}$ by the properties of Delayer's strategy.

Claim 4.6. *If $T_v \neq \emptyset$, then F_v does not contradict any clause of $\varphi \circ g$.*

Proof. Consider some $\sigma \in T_v$, by the remark above, $\hat{\sigma}|_{\mathrm{Cl}(L(F_v))} \in \mathcal{A}$, hence by the conditions of the theorem, $\hat{\sigma}|_{\mathrm{Cl}(L(F_v))}$ does not falsify any clause of φ. Then by Lemma 2.19, F_v does not contradict any clause of $\varphi \circ g$. □

If a and b are distinct leaves of H, then linear systems F_a and F_b contradict each other. Hence, $T_a \cap T_b = \emptyset$. Thus, to prove the theorem, it is sufficient to show that $\sum_{a\in\mathcal{L}} |T_a| \geq 2^{-(l-1)c} \cdot |T_r|$, where $\mathcal{L}$ denotes the set of leaves of H.

For every vertex v in H with children u and w, we define p_v such that $p_v := \frac{|T_w|+|T_u|}{|T_v|}$ if $|T_v| \neq 0$ and $p_v := 1$, otherwise. Notice that if $|T_v| = 0$, then $|T_u| = |T_w| = 0$, hence the equality $p_v|T_v| = |T_u| + |T_v|$ is always satisfied. Since u and w are children of v, there exists a linear form f and $\alpha \in \{0, 1\}$ such that $F_u = F \wedge (f = \alpha)$ and $F_v = F \wedge (f = 1 - \alpha)$. Hence, $\mathrm{Cl}(L(F_u)) = \mathrm{Cl}(L(F_w)) = \mathrm{Cl}(L(F_v) \cup \{f\})$. It is easy to see that $T_u \cup T_v$ is the set of assignments τ satisfying F_v such that $\hat{\tau}|_{\mathrm{Cl}(F_v\cup\{f\})}$ is consistent with Delayer's strategy if variables appear in the order $\theta_u = \theta_w$. By Lemma 4.5 applied to the order θ_u, we get that $p_v \geq 2^{-(\ell-1)k}$, where k is the number of forced variables in $\mathrm{Cl}(L(F_u)) \setminus \mathrm{Cl}(L(F_v))$ according to Delayer's strategy if variables appear in the order θ_u.

Let a be a leaf of H. Consider a path from the root of H to a: $u_1 = r, u_2, \ldots, u_s = a$. As we noticed above, $p_{u_i} \geq 2^{-(\ell-1)k_i}$, where k_i is the number of forced variables in $\mathrm{Cl}(F_{u_{i+1}}) \setminus \mathrm{Cl}(F_{u_i})$ according to Delayer's strategy if variables appear in the order θ_a. By the properties of the strategy, Delayer should spend at most c black coins if he earns at most w. Since $|\mathrm{Cl}(L(F_a))| \leq w + |\rho_0|$, $\prod_{i=1}^{s-1} p_{u_i} \geq 2^{-(\ell-1)c}$.

Recall that $\mathcal{L}$ is the set of leaves of H. By Lemma 4.4 applied to H, p_v and $n_v := |T_v|$, $\sum_{a\in\mathcal{L}} |T_a| \geq |T_r|2^{-(\ell-1)c}$. □

4.4 Proof of Lemma 4.5

Lemma 4.7. *Let $g_1, g_2, \ldots, g_n$ be 2-stifling gadgets from $\{0,1\}^\ell \to \{0,1\}$. Let $\alpha_1, \alpha_2, \ldots, \alpha_n$ be arbitrary Boolean functions from $\{0,1\}^{(\ell-1)n} \to \{0,1\}$ and $\beta_1, \beta_2, \ldots, \beta_n$ be affine functions from $\{0,1\}^{(\ell-1)n} \to \{0,1\}$. Then there exist $r_1, r_2, \ldots, r_n \in \{0,1\}^{\ell-1}$ such that for every $i \in [n]$, $g_i(r_i, \alpha_i(r_1, r_2, \ldots, r_n)) = \beta_i(r_1, r_2, \ldots, r_n)$.*

Proof of Lemma 4.7. We have to prove that the system of n equations has a solution. We introduce variables $r_{i,1}, r_{i,2}, \ldots, r_{i,\ell}$ for bits of r_i. We prove the lemma by induction on the number of equations n.

Assume that there is $i \in [n]$ such that β_i is a constant. Then we can satisfy the ith equation by choosing an appropriate value of r_i by using a 2-stifling property of g. Thus we fix values of variables corresponding to r_i and remove the ith equation. The remaining equations we can satisfy by the induction hypothesis.

Assume that there exists $i \in [n]$ such that $\beta_i(r_1, \ldots, r_n)$ is dependent on $r_{i,j}$ for some $j \in [\ell]$. Then the equation $\beta_i = 0$ expresses $r_{i,j}$ from the other variables. We change all occurrences of $r_{i,j}$ to this expression. Now on the left-hand side, we have g_i applied to $r_{i,k}$ for $k \in [\ell-1] \setminus \{j\}$ and two more complicated positions. After this, we fix values of $r_{i,k}$ for $k \in [\ell-1] \setminus \{j\}$ such that the value of the gadget equals zero regardless of the values of the two remaining positions. This is possible since g_i is 2-stifling. Thus we eliminate variables r_i and delete the ith equation. The remaining equations can be satisfied by the induction hypothesis.

Let us consider a directed graph with vertices $[m]$. We say that there is an edge from i to j if β_i depends on a variable corresponding to r_j. The graph contains a directed cycle since every vertex has an outgoing edge.

Consider the minimal directed cycle: β_{i_1} depends on the variable r_{i_2,j_2}, β_{i_2} depends on r_{i_3,j_3} etc., β_{i_k} depends on r_{i_1,j_1}. Notice that variables $r_{i_1,j_1}, r_{i_2,j_2}, \ldots, r_{i_l,j_k}$ have exactly one occurrence in $\beta_{i_1}, \beta_{i_2}, \ldots, \beta_{i_k}$ since otherwise the cycle can be decreased. Consider the linear system $\beta_{i_1} = 0, \wedge \cdots \wedge \beta_{i_l} = 0$. From this system we can express variables $r_{i_1,j_1}, r_{i_2,j_2}, \ldots, r_{i_l,j_k}$ in other variables using these equations. We substitute these expressions instead of variables. After this substitution the right-hand sides of equations with numbers $i_1, i_2, \ldots, i_k$ will be fixed to 0. For every $j \in [k]$ the left-hand side of i_jth equation contains $\ell-2$ positions that contain only variables, so we can fix the values of all variables to satisfy the i_jth equation using the 2-stifling property. Thus we eliminate all variables corresponding r_i and delete ith equation for $i \in \{i_1, \ldots, i_k\}$. The remaining equations can be satisfied by the induction hypothesis. □

Proof of Lemma 4.5. If $\mathrm{Cl}(L(F)) = \mathrm{Cl}(L(F) \cup \{f\})$, then the lemma is trivial. So we assume that $\mathrm{Cl}(L(F)) \subsetneq \mathrm{Cl}(L(F) \cup \{f\})$. Let ρ be a solution of F restricted to variables with support in $\mathrm{Cl}(F)$ such that $\hat{\rho}$ is consistent with the strategy h if variables appear in the order θ. Let $T_\rho = \{\tau \in T \mid \tau \text{ is consistent with } \rho\}$. Let $T'_\rho = \{\tau \in T' \mid \tau \text{ is consistent with } \rho\}$. It is sufficient to show that for any ρ, $|T'_\rho| \geq |T_\rho| 2^{-(\ell-1)n}$.

The linear system $F|_\rho$ is satisfiable, and the set of its linear forms is safe by the definition of closure. Hence, by Theorem 2.4, one can choose a basis of the span of the columns of the matrix of $F|_\rho$ such that there is at most one basis element in every block. Let Z denote the set of variables corresponding to this basis. Every solution of $F|_\rho$ defines an element of T_ρ. Hence, the size of T_ρ equals the number of solutions of $F|_\rho$. The set of solutions of $F|_\rho$ can be constructed as follows: choose arbitrary values of non-Z variables, and then the values of Z-variables are uniquely determined. Let D denote the number of non-Z variables in $F|_\rho$. Then $|T_\rho| = 2^D$.

Let K be the set of unlifted variables from $\mathrm{Cl}(L(F) \cup \{f\}) \setminus \mathrm{Cl}(L(F))$ that are forced in the strategy h where variables appear in the order θ. We will show that if we arbitrarily fixvalues of non-Z variables such that their support is not in K and values of the $x_{i,\ell}$ if $y_i \in K$ and Z does not contain variables with support i, then we can extend this assignment to an element of T'_ρ. The number of unfixed variables out of Z is exactly $(\ell-1)n$. Hence we will get the desired inequality $|T'_\rho| \geq 2^{D-(\ell-1)n} = |T_\rho| 2^{-n(\ell-1)}$.

For all blocks that are not in K, we have fixed values of variables out of Z. For every $y_i \in K$ we have to choose values of unfixed variables from $x_{i,1}, x_{i,2}, \ldots, x_{i,\ell}$ such that when we determine values of Z variables, the value of the gadget applied to $x_{i,1}, x_{i,2}, \ldots, x_{i,\ell}$ will be consistent with the strategy h.

We know that $|K| = n$; w.l.o.g. $K = \{y_1, y_2, \ldots, y_n\}$. According to the strategy h values of forced variables are computed by linear functions from the values of the previous variables. W.l.o.g. we assume that the correct values of these variables depend only on unforced variables. Unfortunately, it is possible that when we have fixed all lifted variables except one in the block with support not in K, the value of the gadget is not determined and depends on the value of the last variable. Then, the gadget's value is linearly dependent on the unfixed variable from Z. In this case, the required value of some variable from K may depend on these unfixed Z-variables, but in this case, the dependence is linear. For every $y_i \in K$ we denote unfixed variables among $x_{i,1}, x_{i,2}, \ldots, x_{i,\ell}$ by vector of variables r_i; for all $i \in [n]$, r_i consists of exactly $\ell - 1$ variables. Notice that the values of Z-variables are chosen to satisfy the system $F|_\rho$. Hence, the values of every Z-variable can be computed from $r_1, r_2, \ldots, r_\ell$ by an affine function. Thus, we have to satisfy the following system of equations: $\bigwedge_{i=1}^{n} g_i(r_i, \alpha_i(r_1, r_2, \ldots, r_n)) = \beta_i(r_1, r_2, \ldots, r_n)$, where for $i \in [n]$, g_i is a function obtained from g by a variable permutation, α_i and β_i are affine functions; α_i corresponds to either expression of a Z-variable (if there are Z-th variables in ith block) or constant (if there are no Z variables in ith block and we have fixed $x_{i,\ell}$); β_i corresponds the linear dependence between the value of the variable y_i according the strategy h and values of several Z-variables.

This system of equations has a solution by Lemma 4.7. □

5 Lifting from Strategies in Advanced Prover-Delayer Games to Regular Resolution over Parities

In this section, our goal is to show that if a formula Φ has some good properties and $g : \{0,1\}^\ell \to \{0,1\}$ is a 2-stifling gadget, then $\Phi \circ g$ requires large regular Res($\oplus$) proofs (for precise statement see Theorem 5.2). The high-level proof plan is as follows: 1) We consider a random full assignment σ of the variables of $\Phi \circ g$ and make several steps in a branching program associated with a regular Res($\oplus$) refutation of $\Phi \circ g$ from the source according to σ. Let C be a clause at the end of the path. 2) We show that with probability p the linear system $\neg C$ has rank at least r. 3) By the construction σ satisfies $\neg C$. Random assignment satisfies a linear system with rank at least r with probability at least 2^{-r}. Hence, the refutation must contain at least $p2^r$ clauses.

The main technical part is the realization of the step 2 of the above plan. In Subsection 5.1, we define the notion of a q-corect partial assignment of an unlifted formula Φ. In Theorem 5.1, we show that if a clause C is on a large enough distance from the source of the branching program and $\neg C$ has a solution π such that $\hat{\pi}|_{\mathrm{Cl}(L(C))}$ is q-correct, then $\neg C$ has a large rank.

In Subsection 5.2 we define the advanced (Φ, q) games for the formula Φ and the set of q-correct partial assignments as a special case of games defined in Section 4. Using Theorem 4.3, we consider

a random walk described in the 1st step of the above plan and show that if Delayer has a good strategy in the game and get that with high enough probability $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ is q-correct. Then, we prove the main result (Theorem 5.2).

In Subsection 5.3, we give an example of the application of Theorem 5.2 for lifted Tseitin formulas.

5.1 Rank Lower Bound and q-Correct Partial Assignments

Let Φ be an unsatisfiable CNF formula that can be represented in the form of $\bigwedge_{v\in V}\phi_v$, where ϕ_v is a CNF formula, in which each clause consists of the same set of variables. In the simple case, each ϕ_v contains just one clause. For example, for Tseitin formulas, ϕ_v can be a parity condition in the vertex v.

A partial assignment ρ is called *q-correct* for Φ if for every set $U \subseteq V$ such that $|\mathrm{Vars}(\bigwedge_{v\in U}\phi_v)| < |\mathrm{Vars}(\Phi)| - q$, ρ can be extended to an assignment satisfying $\bigwedge_{v\in U}\phi_v$.

Theorem 5.1. *Let $g\colon \{0,1\}^\ell \to \{0,1\}$ be a 1-stifling gadget. Consider a regular* Res($\oplus$) *refutation of the lifted formula $\Phi\circ g$ and its linear branching program. Let C be a linear clause such that there is a path of length t from the source of the linear branching program to C. Suppose that $\neg C$ has a solution σ such that $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ is q-correct (here $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ is the restriction of $\hat{\sigma}$ to $\mathrm{Cl}(L(C))$ which is identified with the set of unlifted variables). Then* $\mathrm{rk}(\neg C) \geq t - \ell q$.

Proof. Consider the linear branching program associated with the Res($\oplus$) refutation of $\Phi\circ g$. Let W consist of all sinks u of the linear branching program such that there is a path from C to u and the conjunction of linear equations labeling the edges of this path is consistent with the linear system $\neg C$ (i.e., the conjunction of the linear system on the path and $\neg C$ is satisfiable). Let A be the set of labels of the nodes from W; A consists of clauses of $\Phi\circ g$. It is easy to see that A semantically implies C. Indeed, consider an assignment σ of the lifted variables that falsifies C. We start a path in the linear branching program from C to a sink such that σ satisfies all equalities along the edges. Let the path end in a sink w labeled with a clause D. By Lemma 2.1, σ falsifies D.

Let $U := \{v \in V \mid \exists C \in A, C \text{ is a clause of } \phi_v \circ g\}$.

Assume that $|\mathrm{Vars}(\bigwedge_{v\in U}\phi_v)| < |\mathrm{Vars}(\Phi)| - q$, then since $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ is q-correct, there exists τ extending $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ such that τ satisfies $\bigwedge_{v\in U}\phi_v$. By Lemma 2.18, there exists a full assignment π to lifted variables such that $\hat{\pi} = \tau$ and π satisfies $\neg C$. Then π satisfies $\bigwedge_{v\in U}\phi_v\circ g$, hence, π satisfies all clauses from A. Since C is a semantic implication of A, π satisfies C, this is a contradiction since π satisfies $\neg C$.

Hence, $|\mathrm{Vars}(\bigwedge_{v\in U}\phi_v)| \geq |\mathrm{Vars}(\Phi)| - q$. Since for all $v \in V$, all clauses from ϕ_v have the same set of variables, Lemma 2.16 implies that $|\mathrm{Vars}(A)| = |\mathrm{Vars}(\bigwedge_{v\in U}\phi_v \circ g)|$. Again by Lemma 2.16, $|\mathrm{Vars}(A)| \geq \ell(|\mathrm{Vars}(\Phi)| - q) \geq |\mathrm{Vars}(\Phi\circ g)| - q\ell$.

Consider $W = \langle L(C) \cup \mathrm{Post}(C)\rangle$. Using regularity, by Lemma 2.2, we get that $\dim(\mathrm{Post}(C)) \leq |\mathrm{Vars}(\Phi\circ g)| - t$, and thus $\dim(W) \leq \dim\langle L(C)\rangle + \dim(\mathrm{Post}(C)) \leq \mathrm{rk}(\neg C) + |\mathrm{Vars}(\Phi\circ g)| - t$. On the other hand, for every clause $D \in A$, there is a path from C to D such that $\neg C$ is consistent with the system of all equations labeling the path's edges. By Lemma 2.1, all variables that appear in D are linear combinations of $L(C)$ and the linear forms of the equations at the edges of this path from C to D.

Hence, $\dim(W) \geq |\mathrm{Vars}(A)| \geq |\mathrm{Vars}(\Phi\circ g)| - q\ell$. Combining those two inequalities together, we get $\mathrm{rk}(\neg C) \geq t - q\ell$. □

5.2 Lower Bound for Regular Res($\oplus$)

Let Φ be an unsatisfiable CNF formula that can be represented in the form of $\bigwedge_{v\in V}\phi_v$, where ϕ_v is a CNF formula, in which each clause consists of the same set of variables. Let $\mathcal{A}_q$ be the set of all q-correct for Φ partial assignments. By *advanced* (Φ,q)*-game of Prover and Delayer* we denote $(\Phi,\mathcal{A}_q)$ games with empty starting position.

Theorem 5.2. *Let Φ be an unsatisfiable CNF formula. Assume that in the (Φ,q)-game, Delayer has a linearly described strategy that guarantees him to earn t white coins while paying at most c black coins, where $t+q<|\mathrm{Vars}(\Phi)|$. Let $g:\{0,1\}^\ell\to\{0,1\}$ be 2-stifling gadget. Then the size of any regular* Res($\oplus$) *refutation of $\Phi\circ g$ is at least $2^{t-q\ell-c(\ell-1)}$.*

Proof. Consider a regular Res($\oplus$) refutation of $\Phi\circ g$ and the linear branching program associated with it. Consider a random full assignment σ of variables $\Phi\circ g$ and make t steps according to σ in the linear branching program starting from the source. The condition $t+q<|\mathrm{Vars}(\Phi)|$ implies that any q-correct assignment of size at most t does not contradict any clause of Φ. Thus, by Theorem 4.3 with probability $2^{-c(\ell-1)}$ we reach a node labeled with a linear clause C such that the partial assignment $\hat{\sigma}|_{\mathrm{Cl}(L(C))}$ is q-correct. Then by Theorem 5.1, $\mathrm{rk}(\neg C)\geq t-q\ell$. Hence, a random assignment satisfies $\neg C$ with probability at most $2^{-t+q\ell}$. Thus, the refutation consists of at least $2^{t-q\ell-c(\ell-1)}$ linear clauses. □

5.3 Lifted Tseitin formulas are hard for regular Res($\oplus$)

In this subsection, we give an example of the application of Theorem 5.2, namely, we show that lifted Tseitin formulas based on spectral expanders are hard for regular Res($\oplus$).

Theorem 5.3. *Let $\mathrm{T}(G,c)$ be an unsatisfiable Tseitin formula based on a spectral (n,d,α)-expander $G(V,E)$, where $\alpha<1/2$ and $d\geq 4$. Let $\beta\leq\frac{1}{4}$ and $t=\beta n$ be a natural number. Then in the advanced $(\mathrm{T}(G,c),\epsilon t)$-game, the natural strategy of Delayer (defined in Section 4.2) guarantees him to earn t white coins while paying at most $\frac{2}{d(1-\alpha)}t$ black coins, where $\epsilon=\frac{\alpha}{1-\alpha}+\frac{2\beta}{d}\cdot\frac{1}{(1-\alpha)^2}$.*

Proof. Since $\alpha<1/2$ and $d\geq 4$, $t=n/4<\frac{dn}{4}\cdot\frac{1-\alpha}{2}$.

Consider $\rho\in\mathcal{T}$ such that $|\rho|\leq t$. Let us show that ρ is ϵt-correct.

Let G_ρ be the graph of the Tseitin formula $\mathrm{T}(G,c)|_\rho$. Let us denote by A the set of all vertices that do not belong to the maximal connected component of G_ρ. Since $\rho\in\mathcal{T}$ and $|\rho|\leq t<\frac{dn}{4}\cdot\frac{1-\alpha}{2}$, $|A|<n/2$. Since every move in the game corresponds to removing the edge in the graph, during the first t steps we have removed at most t edges from G. Hence, by Lemma 2.14, $t\geq E(A,V\setminus A)\geq d|A|\frac{1-\alpha}{2}$. Thus, $|A|\leq\frac{2t}{d(1-\alpha)}$.

Notice that $|E(A,A)|$ equals two times the number of edges in G with both ends inside A. Using Lemma 2.14, we can estimate the number of edges inside A as follows: $\frac{1}{2}|E(A,A)|\leq\frac{1}{2}\left(\alpha d|A|+\frac{d|A|^2}{n}\right)=\frac{1}{2}|A|d(\alpha+|A|/n)\leq\frac{t}{1-\alpha}\left(\alpha+\frac{2}{d(1-\alpha)\frac{t}{n}}\right)=\beta n\left(\frac{\alpha}{1-\alpha}+\frac{2\beta}{d}\cdot\frac{1}{(1-\alpha)^2}\right)=\epsilon t$.

Let $P_\rho(v)$ denote the parity condition of the Tseitin formula $\mathrm{T}(G,c)|_\rho$ for vertex $v\in V$. Consider a inclusion minimal set $U\subseteq V$ such that $\bigwedge_{v\in U}P_\rho(v)$ is unsatisfiable. Since $\rho\in\mathcal{T}$, the minimality of U implies that U contains only vertices from the largest connected component. By Corollary 2.10, U contains all vertices from the largest component. Hence, $|\mathrm{Vars}(\bigwedge_{v\in U}P_\rho(v))|=|\mathrm{Vars}(\mathrm{T}(G,c))|-\frac{1}{2}|E(A,A)|\geq|\mathrm{Vars}(\mathrm{T}(G,c))|-\epsilon t$.

Whenever we remove a bridge, the number of connected components increases by one. Lemma 2.15 implies that during t moves, the number of times when the requested edge is a bridge is at most $\frac{2}{d(1-\alpha)}t$. Hence, hence after t moves, Delayer, using the natural strategy, earns t white coins and pays at most $\frac{2}{d(1-\alpha)}t$ black coins. □

Corollary 5.4. *Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget and G be an (n, d, α)-expander, where $d \geq 6\ell$ and $\alpha \leq \frac{1}{6\ell}$. Then, the size of any regular* Res(⊕) *refutation of* $\mathrm{T}(G, c) \circ g$ *is at least* $2^{\Omega(n)}$.

Proof. Let $\beta = \frac{\lfloor n/4 \rfloor}{n}$, clear that $\beta = 1/4$. By Theorem 5.3, Lemma 4.2 and Theorem 5.2, size of any regular Res(⊕) refutation of $\mathrm{T}(G, c) \circ g$ is at least $2^{t(1-\epsilon\ell)-t\left(\frac{2}{d(1-\alpha)}\right)(\ell-1)}$, where $\epsilon = \frac{\alpha}{1-\alpha} + \frac{2\beta}{d} \cdot \frac{1}{(1-\alpha)^2}$. It is sufficient to have $\epsilon + \frac{2}{d(1-\alpha)} = \frac{\alpha}{1-\alpha} + \frac{2\beta}{d} \cdot \frac{1}{(1-\alpha)^2} + \frac{2}{d(1-\alpha)} < 1/\ell$.

It is easy to verify that this is true since $d \geq 6\ell$, $\beta \leq \frac{1}{4}$, and $\alpha \leq \frac{1}{6\ell}$:

$$\frac{\alpha}{1-\alpha} + \frac{2\beta}{d} \cdot \frac{1}{(1-\alpha)^2} + \frac{2}{d(1-\alpha)} \leq \frac{1}{6\ell} \cdot \frac{6}{5} + \frac{1}{6\ell} \cdot \frac{1}{2} \cdot \left(\frac{6}{5}\right)^2 + \frac{1}{3\ell} \cdot \frac{6}{5} \leq \frac{1}{5\ell} + \frac{1}{6\ell} + \frac{2}{5\ell} < 1/\ell.$$

□

6 Lifting from Resolution Depth

In this section, we show how to construct formulas that require large regular Res(⊕) refutations based on formulas requiring large resolution depth. In Subsection 6.1, we define simplified games that are very similar to games characterizing resolution depth. We show that the strategy in these simplified games can be converted into the strategy in advanced games for the formula lifted by the parity gadget. In Subsection 6.3, we define a mixing transformation of formulas that does not change the formula semantically but allows us to convert strategies in the games characterizing depth to the simplified games.

6.1 Simplified Games

Let Φ be an unsatisfiable CNF formula that can be represented in the form of $\bigwedge_{v \in V} \phi_v$, where ϕ_v is a CNF formula, in which each clause consists of the same set of variables. We define one more game associated with Φ and a natural number q.

A simplified (Φ, q)-game of Prover and Adversary. In this game there are two players: Prover and Adversary. On every move, Prover chooses a variable x of the formula Φ, and Adversary chooses the 0/1 value of this variable. The game ends when the current partial assignment is not q-correct. For every his move Adversary earns a coin.

A *strategy* for the Adversary is a function $f \colon H \times \mathrm{Vars}(\Phi) \to \{0,1\}$, where H is the set of all possible sequences of k queries asked by Delayer in the previous rounds (so it is the sequences of the form $(x_{i_1}, x_{i_2}, \ldots, x_{i_k})$). The strategy is utilized naturally: on every step, given a previous sequence of queries Q of Delayer and a last queried variable x, Adversary answers $f(Q, x)$. Notice that in this definition f *does not depend* on the previous answers of Delayer since Delayer can compute these answers by itself if necessary.

If $g : \{0,1\}^\ell \to \{0,1\}$ is some gadget, then $\Phi \circ g$ can be represented as $\bigwedge_{v \in V}(\phi_v \circ g)$. By Lemma 2.16, all clauses of $\phi_v \circ g$ use the same set of variables.

Lemma 6.1. *Assume that there is a strategy of Adversary in the simplified (Φ, q)-game that allows him to earn at least t coins. Let $\oplus_r : \{0,1\}^r \to \{0,1\}$ be the parity function. Then for the advanced $(\Phi \circ \oplus_r, qr)$-game, there is a linearly described strategy of Delayer that guarantees him to earn tr white coins while paying at most t black coins.*

Proof. Let Φ depend on variables $y_1, y_2, \ldots, y_m$. Then $\Phi \circ \oplus_r$ depends on variables $\{x_{i,j} \mid r \in [m], j \in [r]\}$.

Let us present a linear description f of Delayer's strategy. Let L be an ordered list of already asked variables and $x_{i,j}$ be a new variable.

- If $|L \cap \{x_{i,1}, x_{i,2}, \ldots, x_{i,r}\}| < r - 1$, then $f(L, x_{i,j}) = *$.
- Otherwise, let $L' = \{y_j \mid \{x_{j,1}, x_{j,2}, \ldots, x_{j,r}\} \subseteq L\}$. Let for $y_i \in L$, n_i be the maximal number in the list L of a variable with support i. We introduce an order in L': y_k is less than y_j if $n_k < n_j$. Consider the strategy of Adversary in the simplified game that guarantees him to earn t coins. Assume that Prover chooses variables according to the introduced order in L' and then asks y_i. Let $\alpha \in \{0,1\}$ be the value of y_i according to this strategy. In this case $f(L, x_{i,j}) := \sum_{k \in [r] \setminus j} x_{i,k} + \alpha$.

Let us show that this strategy guarantees Delayer to earn tr white coins. Consider any subset $U \subseteq V$ such that $|\mathrm{Vars}(\bigvee_{u \in U} \phi_u \circ \oplus_r)| < |\mathrm{Vars}(\Phi \circ \oplus_r)| - qr$. By Lemma 2.16, $|\mathrm{Vars}(\bigvee_{u \in U} \phi_U)| < |\mathrm{Vars}(\Phi)| - q$. Assume that Delayer uses the described strategy. Consider some state of the game after at most tr moves. Let $M = \{k \in [m] \mid x_{k,j}$ were requested for all j $\in [r]\}$; $|M| \leq t$. There are at most t such k such that $x_{k,i}$ were asked. For $m \in [k]$, values of $\sum_{i=1}^r x_{k_i}$ are fixed according to the value of y_k to the strategy of Adversary in the simplified game. Hence the current partial assignment in the variables $\{y_k \mid k \in M\}$ can be extended to satisfy $\bigvee_{u \in U} \phi_U$. So we can extend partial assignment in the lifted variables to satisfy $\bigvee_{u \in U}(\phi_u \circ \oplus_r)$.

It is easy to see that for every moment when Delayer pays a black coin, there are at least $(r-1)$ moments when he does not pay. Hence the number of paid black coins is at most t. □

6.2 Resolution Depth

Simplified games are similar to games characterizing resolution depth [30].

The depth of a resolution proof is the length of the shortest path between an empty clause and a clause of the refuted formula. The resolution depth of an unsatisfiable CNF formula φ denoted by $d_R(\varphi)$ is the minimal possible depth overall resolution refutations of φ. The resolution depth of an unsatisfiable CNF formula φ can be characterized by the following game of Prover and Adversary: on every move, Prover asks the value of the variable of φ and Adversary answers and earns a coin. The game ends whenever the current assignment falsifies the clause of φ.

Lemma 6.2 ([30])**.** *$d_R(\varphi) \geq t$ iff Adversary has a strategy that guarantees him to earn at least t coins.*

The main difference between simplified (Φ, q)-games and games characterizing depth is the condition of the end of the game. In the next subsection, we define the mixing operation for formulas to map formulas with large resolution depth to formulas with good Adversary strategies in simplified games.

6.3 Mixed Formulas

6.3.1 Mixers

A bipartite graph $G(X, Y, E)$ is an (n, D, α, ϵ)-mixer if

1. $|X| = n$, $|Y| = n$;
2. Degrees of all vertices from X are at most D;
3. For every $A \subseteq X$ and $B \subseteq Y$ if $|A| \geq \alpha n$ and $|B| \geq \epsilon n$, then there is at least one egde between A and B.

Lemma 6.3. *For every integer n, real $\alpha \in (0; 1)$ and $\epsilon \in (0, 1)$ there exists an (n, D, α, ϵ)-mixer, where $D = O(\frac{1}{\epsilon\alpha})$.*

Proof. We construct a graph with vertices X and Y such that $|X| = n, |Y| = n$ by the random process.

1. Initially, the set of edges E is empty;
2. For every $v \in X$ repeat $D = \lceil K\frac{1}{\alpha\epsilon}\rceil$ times (the value of K will be chosen later):
    3. Choose $u \in Y$ at random;
    4. Add (v, u) to E;

For every $A \subseteq X$ and $B \subseteq Y$ such that $|A| \geq \alpha n$, $|B| \geq \epsilon n$, the probability that there are no edges between A and B is at most $(1 - \epsilon)^{D|A|} \leq (1 - \epsilon)^{Kn/\epsilon} < e^{-Kn}$.

The number of pairs A and B is at most 2^{2n}. Hence by the union bound the probability that there exist such $A \subseteq X$ and $B \subseteq Y$ such that $|A| \geq \alpha n$, $|B| \geq \epsilon n$ and there are no edges between A and B is at most $e^{-Kn}2^{2n}$ that is less than 1 for $K \geq 2$. Hence with positive probability, the constructed graph is an (n, D, α, ϵ)-mixer with $D = O(\frac{1}{\alpha\epsilon})$. □

The explicit constructions of mixers can also be obtained from spectral expanders using the expander mixing lemma (Lemma 2.14).

6.3.2 Mixed formulas

Let C be a clause and Z be a set of propositional variables with no occurrences in C. Let Clauses(Z) be the set of all $2^{|Z|}$ different clauses, each containing all variables from Z. We denote by pad(C, Z) the CNF formula $\bigwedge_{D\in\text{Clauses}(Z)}(C \vee D)$.

Lemma 6.4. *There is a resolution derivation of C from* pad(C, Z) *of length $2^{|Z|}$ and of depth $|Z|$.*

Proof. The proof is straightforward by induction on the number of variables in Z. □

Lemma 6.5. *C is semantically equivalent to* pad(C, Z).

Proof. pad$(C, Z) = \bigwedge_{D\in\text{Clauses}(Z)}(C \vee D)$ and it is semantically equivalent to $C \vee \bigwedge_{D\in\text{Clauses}(Z)} D$ and the later is semantically equivalent to C since $\bigwedge_{D\in\text{Clauses}(Z)} D$ is identically false. □

Let $\varphi = \bigwedge_{v \in V} C_v$ be a CNF formula from n variables (for every $v \in V$, C_v is a clause) and $G(X, Y, E)$ be a bipartite graph with $|X| = |Y| = n$ and with degrees of all vertices from X at most D. We define a CNF formula $\mathrm{mix}_G(\varphi)$ as follows:

- Let π_1 be a bijection from $\mathrm{Vars}(\varphi) \to X$ and π_2 be a bijection from $Y \to \mathrm{Vars}(\varphi)$.
- $\mathrm{mix}_G(\varphi) = \bigwedge_{v \in V} \psi_v$, where $\psi_v = \mathrm{pad}(C_v, \pi_2(\Gamma(\pi_1(\mathrm{Vars}(C_v)))) \setminus \mathrm{Vars}(C_v))$, where for $A \subseteq X$, $\Gamma(A)$ is the set of neighbors of the set A in the graph G.

Notice that if φ is a k-CNF formula, then $\mathrm{mix}_G(\varphi)$ is a kD-CNF formula. By Lemma 6.5, $\mathrm{mix}_G(\varphi)$ is semantically equivalent to φ.

Lemma 6.6. *If G is (n, D, α, ϵ)-mixer, then if $\mathrm{Vars}(\bigwedge_{v \in V} C_v) \geq \alpha n$, then $\mathrm{Vars}(\bigwedge_{v \in V} \psi_v) \geq (1 - \epsilon)n$.*

Proof. The proof is straightforward. □

6.3.3 Lifting from resolution depth

Lemma 6.7. *Let a CNF formula φ with n variables have a resolution depth at least d; let G be $(n, D, \frac{d}{2n}, \epsilon)$-mixer. Then, in the simplified $(\mathrm{mix}_G(\varphi), \epsilon n)$-game, Adversary has a strategy that guarantees him to earn at least $\lfloor d/2 \rfloor$ coins.*

Proof. Adversary will use his strategy in the game characterizing the resolution depth of the formula ϕ, given by Lemma 6.2. Consider the game after $\lfloor d/2 \rfloor$ moves. Let ρ be the current substitution.

Let $\mathrm{mix}_G(\varphi) = \bigwedge_{v \in V} \psi_v$. Assume that for some $U \subseteq V$, ρ can not be extended to satisfy $\bigwedge_{v \in U} \psi_v$. By Lemma 6.5, $\bigwedge_{v \in U} \psi_v$ is semantically equivalent to $\bigwedge_{v \in U} C_v$. Since ρ is the first part of the strategy in the game characterizing depth, $\mathrm{Vars}(\bigwedge_{v \in U} C_v) \geq \frac{d}{2}$ (otherwise, Prover can just query all the variables from $\mathrm{Vars}(\bigwedge_{v \in U} C_v)$ and end the game in less than $\frac{d}{2}$ steps). Hence, by Lemma 6.6, $\mathrm{Vars}(\bigwedge_{v \in U} \psi_v) \geq (1 - \epsilon)n$. □

Theorem 6.8. *Let φ_n be the family of unsatisfiable k-CNF formulas in n variables such that $d_R(\varphi_n) \geq \alpha n$. Let G be a (n, D, α, ϵ)-mixer, where $\epsilon = \alpha/100$, $D = O(\frac{1}{\alpha^2})$, that exists by Lemma 6.3. Then any regular $\mathrm{Res}(\oplus)$ refutation of $\mathrm{mix}_G(\varphi_n) \circ \oplus_5 \circ Maj_5$ has size at least $2^{\alpha n/4 - 1}$.*

Notice that if in the conditions of Theorem 6.8, k and α are constants, then $\mathrm{mix}_G(\varphi_n) \circ \oplus_5 \circ Maj_5$ is $O(k)$-CNF formula.

Proof. By Lemma 6.7, in the simplified $(\mathrm{mix}_G(\varphi_n), \epsilon n)$-game, there is a strategy of Adversary that guarantees him to earn at least $\lfloor \alpha n/2 \rfloor$ coins.

By Lemma 6.1, in the advanced $(\mathrm{mix}_G(\varphi_n) \circ \oplus_5, 5\epsilon n)$ game there is a strategy of Delayer that guarantees him to earn at least $5\lfloor \alpha n/2 \rfloor$ white coins while paying at most $\lfloor \alpha n/2 \rfloor$ black coins.

$Maj_5 : \{0,1\}^5 \to \{0,1\}$ is a 2-stifling gadget; hence, by Theorem 5.2, the size of any regular $\mathrm{Res}(\oplus)$ refutation of $\mathrm{mix}_G(\varphi_n) \circ \oplus_5 \circ Maj_5$ is at least $2^{5\lfloor \alpha n/2 \rfloor - 25\epsilon n - 4\lfloor \alpha n/2 \rfloor} \geq 2^{\alpha n/4 - 1}$. □

7 Regular Res($\oplus$) Does Not Simulate Resolution

In this section, we give an alternative and improved separation between regular Res($\oplus$) and Resolution firstly proved by Bhattacharya, Chattopadhyay, and Dvorak [7].

One of the possible ways to do it using our technique is to apply the combination of mixing and lifting from the previous section to pebbling formulas Peb(G_n) that have $O(n)$ variables and resolution depth $\Omega(n/\log n)$ [30, 26]. The problem is that we need $(n, D, O(1/\log n), O(1/\log n))$ mixers, and for them, $D = O(\log^2 n)$ and the resulting formula will have superpolynomial size. This will imply some separation but not very good. Instead, we will consider Pebbling formulas on the well-structured grid graphs. For such formulas, we can require a much weaker mixing property that allows us to decrease the degree of mixers to $O(\log n)$.

Let $H_n(V_n, E_n)$ be a directed grid graph with the set of vertices $V_n = [n] \times [n]$. The edges are oriented to the left and the bottom or formally 1) for $i > 1, j > 1$ the vertex (i, j) has two outgoing edges to $(i-1, j)$ and $(i, j-1)$; 2) for $j > 1$, the vertex $(1, j)$ has one outgoing edge to $(1, j-1)$; and for $i > 1$ the vertex $(i, 1)$ has one outgoing edge to $(i-1, 1)$.

We define the Pebbling formula Peb(H_n) as follows. The set of variables is $\{x_v \mid v \in V_n\}$. The formula Peb(H_n) is defined to be $\neg x_{1,1} \wedge \bigwedge_{v \in V_n} C_v$, where $C_v = x_v \vee \bigvee_{u \in V_n : (u,v) \in E_n} \neg x_u$.

A bipartite graph $G(X, Y, E)$ is called (n^2, D, ϵ)-grid mixer if 1) $|X| = |Y| = n^2$; let σ be a bijection from $V_n \to X$; 2) degree of all vertices from X are at most D; 3) for all $A \subseteq [n]$ and $B \subseteq [n]$ such that $|A| \geq n/2$ and $|B| \geq n/2$ and all $C \subseteq Y$ such that $|C| \geq \epsilon n$, there is at least one edge between $\sigma(A \times B)$ and C.

Lemma 7.1. *For every integer n and real $\epsilon \in (0, 1)$ there exists an (n^2, D, ϵ)-grid mixer, where $D = O(\log n/\epsilon)$.*

Proof. We construct a graph with vertices X and Y such that $|X| = n^2, |Y| = n^2$ by the random process. Let σ be a bijection from $[n] \times [n] \to X$.

1. Initially, the set of edges E is empty;
2. For every $v \in X$ repeat $D = \lceil K \log n \rceil$ times (the value of K will be chosen later):
    3. Choose $u \in Y$ at random;
    4. Add (v, u) to E;

For every $A \subseteq [n]$ and $B \subseteq [n]$ such that $|A| \geq n/2$, $|B| \geq n/2$ and every $C \subseteq Y$ such that $|C| \geq \epsilon n$, the probability that there are no edges between $\sigma(A \times B)$ and C is at most $(1 - \frac{\epsilon}{n})^{Dn^2/4} \leq (1 - \frac{\epsilon}{n})^{\frac{n}{\epsilon} \cdot \epsilon K n \log n/4} < e^{-K \epsilon n \log n/4}$.

The number of pairs A and B is at most 2^{2n}, and the number of different C is at most $\binom{n^2}{\epsilon n} \leq 2^{2\epsilon n \log n}$. Hence by the union bound the probability that there exist such $A \subseteq [n]$ and $B \subseteq [n]$ and $C \subseteq Y$ such that $|A| \geq n/2$, $|B| \geq n/2$, $|C| \geq \epsilon n$ and there are no edges between $\sigma(A \times B)$ and C is at most $e^{-K n \epsilon \log n/4} 2^{2n + 2n \log n \epsilon}$ that is less than 1 for $K \geq \frac{16}{\epsilon}$. Hence with positive probability, the constructed graph is an (n, D, α, ϵ)-mixer with $D = O\left(\frac{\log n}{\epsilon}\right)$. □

Let $G_{n,\epsilon}$ be an (n^2, D, ϵ)-grid mixer with $D = O\left(\frac{\log(n)}{\epsilon}\right)$. Consider the Pebbling formula $\mathrm{Peb}(H_n) = \neg x_{1,1} \wedge \bigwedge_{v \in V_n} C_v$ and let $\Phi_{n,\epsilon} := \neg x_{1,1} \wedge \mathrm{mix}_G \left(\bigwedge_{v \in V_n} C_v \right)$.

Theorem 7.2. *In the simplified $(\Phi_{n,\epsilon}, \epsilon n)$-game there is a strategy of Adversary that guarantees him to earn at least $n/4$ coins.*

Proof. For every $i \in [n]$ we define the ith *cross* as the set of vertices $\{(a,b) \in V_n \mid a = i \text{ or } b = i\}$.

The top-right part of the ith cross is the set of vertices $\{(a,i),(i,a) \mid a > i\}$.

Let us describe the Adversary's strategy. Adversary has two variables i and p, where i takes values from $[n]$ and p denotes a path in H_n from (i,i) to $(1,1)$. Initially $i = 1$ and p consists of the only vertex $(1,1)$. During the game, Adversary maintains the following invariant:

- For every $i' < i$ there were requests to variables from i'th cross.
- There were no requests to variables from the top-right part of the ith cross.
- If a variable x_v was requested, then x_v was assigned to 0 if v belongs the path p and to 1 otherwise.

The strategy of Adversary is as follows. Let Prover ask the value of x_v

- Adversary responds 0 if v belongs p and 1, otherwise;
- If v belongs to top-right part of the i-th cross, then
  - Let j be the number of the minimal cross such that there were no requests to its variables. If there are no such crosses, Adversary gives up. Notice that the invariant guarantees $j > i$.
  - Since v is the first request to the top-right part of the ith cross, one of the following paths does not contain any requests $(j,j),(j-1,j),\ldots,(i,j),(i,j-1),\ldots,(i,i)$ or $(j,j),(j,j-1),\ldots,(j,i),(j-1,i),\ldots,(i,i)$; let p' denotes this path.
  - $i := j$; the new value of p is p' prolonged by the previous value of p.

Let t be an integer number and $t \le n/4$. Consider the moment after t rounds of the game where Adversary follows the described strategy. Notice that every two crosses have exactly two common vertices, hence there are crosses without requests, and Adversary does not give up. Let k be the number of requests made to the first $i-1$ crosses during the first t rounds. Since every two crosses have two common vertices and all crosses with numbers lesser than i contain requests, $k \ge \frac{i-1}{2}$.

Let $A = \{l \mid i \le l \le n$ and there are no requests to variables $x_{l,j}$ for $j \in [n]\}$ and $B = \{l \mid i \le l \le n$ and there are no requests to variables $x_{j,l}$ for $j \in [n]\}$.

$|A| \ge (n-(i-1)) - (t-k) \ge (n-2k) - (t-k) \ge n - 2t \ge n/2$. Analogously, $|B| \ge n/2$.

Let us verify that the current partial assignment is ϵn-correct.

Let $\Phi_{n,\epsilon} = \neg x_{1,1} \wedge \bigwedge\limits_{v \in V_n} \psi_v$, where ψ_v is the result of pad applied to C_v.

Consider some $U \subseteq V_n$. If $A \times B \subseteq U$, then $\mathrm{Vars}(\bigwedge_{v \in U} \psi_v) \ge n^2 - \epsilon n$ by the property of the greed-mixer G. Assume that there is u such that $u \in A \times B$ and $u \notin U$. We will show that the current assignment can be extended to satisfy $\neg x_{1,1} \wedge \bigwedge\limits_{v \in V_n \setminus \{u\}} C_v$. Since C_v and ψ_v are semantically equivalent, we will get that the current assignment can be extended to satisfy $\neg x_{1,1} \wedge \bigwedge\limits_{v \in U} \psi_v$.

Let $u \in A \times B$. We claim that the formula $\neg x_{1,1} \wedge \bigwedge\limits_{v \in V_n \setminus \{u\}} C_v$ can be satisfied by extending the current assignment from the game. Indeed, there is the following path p' from u to (i,i): at

first, we decrease the first coordinate to level i and then decrease the second coordinate to level i. Consider an assignment that assigns 0 to vertices of p and p' and 1 to all other variables. By the construction of the strategy, this assignment agrees with the current assignment. It is easy to see that C_u is the only unsatisfied condition from $\mathrm{Peb}(H_n)$. Thus, the current assignment is (ϵn)-correct. □

Corollary 7.3. *The size of any regular* $\mathrm{Res}(\oplus)$ *refutation of* $\Phi_{n,\epsilon} \circ \oplus_r \circ \mathrm{Maj}_\ell$ *is at least* $2^{n/4}$*, where* $\ell = 5, r = 6$ *and* $\epsilon = 1/120$.

Proof. By Theorem 7.2 and Lemma 6.1, in the advanced $(\Phi_{n,\epsilon} \circ \oplus_r, \epsilon rn)$ game there is a strategy of Delayer that guarantees him to earn at least $rn/4$ white coins while paying at most $n/4$ black coins.

Since Maj_l is a 2-stifling gadget for $l \geq 5$, by Theorem 5.2, the size of any regular $\mathrm{Res}(\oplus)$ refutation of $\Phi_n \circ \oplus_r \circ \mathrm{Maj}_\ell$ is at least $2^{\frac{n}{4}r - \epsilon r \ell n - \frac{n}{4}(\ell-1)}$. Let us choose $\ell = 5, r = 6$ and $\epsilon = \frac{1}{4r\ell} = 1/120$, then we get the size lower bound $2^{n/4}$. □

Lemma 7.4. *Let* $g : \{0,1\}^r \to \{0,1\}$ *be a gadget. If a CNF formula* φ *has a resolution refutation* Π *of size* S *and width* w *and depth* d*, then the formula* $\varphi \circ g$ *has a resolution refutation of size* $S2^{O(wr)}$ *and depth* $O(dr)$.

Proof. It is enough to show that each step of the resolution derivation Π can be simulated with the lifted derivation of size $2^{O(rw)}$ and depth $O(r)$. Let the set $Y = \{y_1, y_2, \ldots, y_m\}$ be a set of *unlifted* variables and the set $X = \{x_{i,j} \mid i \in [m], j \in [r]\}$ be a set of *lifted* variables.

Consider a resolution step in which we derive $A \vee B$ from $A \vee y_i$ and $B \vee \neg y_i$, where the width of the abovementioned clauses is bounded by w. First, we derive $(A \vee B \vee y_i) \circ g$ from $(A \vee y_i) \circ g$ and $(A \vee B \vee \neg y_i) \circ g$ from $(B \vee \neg y_i) \circ g$ by using the weakening rule. This can be done with a depth 1 parallel derivation with $2^{O(rw)}$ steps. Then both $(A \vee B \vee y_i) \circ g$ and $(A \vee B \vee \neg y_i) \circ g$ consist of at most $2^{O(wr)}$ clauses of the following form: $D \vee D_{y_i}$, where clause D belongs to $(A \vee B) \circ g$ and clause D_{y_i} belongs to $y_i \circ g$ (or $\neg y_i \circ g$).

Now, consider any particular D from $(A \vee B) \circ g$. For any possible subset of literals $l_{i,1}, \ldots, l_{i,r}$ where each $l_{i,j}$ is either $x_{i,j}$ or $\neg x_{i,j}$ we know that $l_{i,1} \vee \ldots \vee l_{i,r}$ belongs as a clause to the CNF corresponding to either $y_i \circ g$ or $(\neg y_i) \circ g$. Thus, D can be derived with the derivation of depth r and size 2^r from $D \vee (y_i \circ g)$ and $D \vee ((\neg y_i) \circ g)$. By applying those derivations in parallel, we get that $(A \vee B) \circ g$ can be derived with size $2^{O(wr)}$ and depth $O(r)$. □

Theorem 7.5. *Suppose* ε, r *and* ℓ *are constants. Then formula* $\Phi_{n,\varepsilon} \circ \oplus_r \circ \mathrm{Maj}_\ell$ *has a resolution refutation of size* $\mathrm{poly}(n)$ *and depth* $O(n)$.

Proof. We consider the following resolution refutation of $\Phi_{n,\varepsilon}$, which consists of two steps:

1. For each $v \in V_n$ we derive ψ_v from $\mathrm{mix}_G(\psi_v)$ by Lemma 6.4. This can be done in parallel with at most $\mathrm{poly}(n)$ steps and both depth and width $O(\log n)$.

2. After deriving ψ_v for each $v \in V_n$ we consider a topological order on the graph H_n. Starting from variable $x_{n,n}$, we derive $x_{n-1,n}$ from $x_{n-1,n} \vee \neg x_{n,n}$ and $x_{n,n-1}$ from $x_{n,n-1} \vee \neg x_{n,n}$. We continue this procedure for each $v \in V_n$ by deriving x_v from $x_v \vee \bigvee_{u \in V_n : (u,v) \in E_n} \neg x_u$ and x_u, where $u \in V_n$ is such that $(u, v) \in E_n$. Since those steps essentially follow the graph H_n, the graph of this refutation has depth $O(n)$ and has $\mathrm{poly}(n)$ steps. In the end we resolve $x_{1,1}$ with $\neg x_{1,1}$ to get an empty clause. The width of this part of the refutation is constant.

All together this gives a refutation of size $\mathrm{poly}(n)$, width $O(\log n)$, and depth $O(n)$. After application of Lemma 7.4 to this refutation with gadget $\oplus_r \circ \mathrm{Maj}_\ell$, we get a resolution refutation of $\Phi_{n,\varepsilon} \circ \oplus_r \circ \mathrm{Maj}_\ell$ of size $\mathrm{poly}(n)$ and depth $O(n)$.

□

8 Size vs Depth Tradeoff for Res(⊕)

Theorem 8.1. *Let φ be an unsatisfiable CNF formula. Let $\mathcal{A}$ be a set of partial assignments for $\mathrm{Vars}(\varphi)$ such that for any $\sigma \in \mathcal{A}$, σ does not falsify any clause of φ and $\mathcal{A}$ is closed under restrictions (i.e., if $\rho \in \mathcal{A}$ and $\tau \subseteq \rho$, then $\tau \in \mathcal{A}$). Assume that there are integers t and c such that for every $\rho \in \mathcal{A}$ such that $|\rho| < t$, in the $(\varphi, \mathcal{A})$-game with start position ρ there is a linearly described strategy of Delayer that guarantees him to earn at least $t - |\rho|$ white coins while paying at most c black coins. Let $g : \{0,1\}^\ell \to \{0,1\}$ be a 2-stifling gadget. Then any $\mathrm{Res}(\oplus)$ refutation of $\varphi \circ g$ has either size at least 2^c or depth at least $\frac{t}{2} \log_{\ell+2}\left(\frac{t}{2c}\right)$.*

Proof. We say that a linear clause C in lifted variables is $\mathcal{A}$-good if there is a solution τ of $\neg C$ such that $\hat{\tau}|_{\mathrm{Cl}(L(C))} \in \mathcal{A}$.

Consider a $\mathrm{Res}(\oplus)$ refutation of $\varphi \circ g$ and denote is by Π.

Claim 8.2. *Assume that Π contains an $\mathcal{A}$-good linear clause C_0 such that $\mathrm{rk}(\neg C_0) \le r$, where $r < t$. Let $S_{t-r}(C_0)$ denote the set of all $\mathcal{A}$-good clauses C such that there is a path from C_0 to C of length $t - r$ in the branching program associated with Π. Assume that for every $C \in S_{t-r}(C_0)$, $\mathrm{rk}(\neg C) \ge r(\ell+1) + c\ell$ holds. Then, the size of the refutation Π is at least 2^c.*

Proof. Since C_0 is $\mathcal{A}$-good, there is a solution τ_0 of $\neg C_0$ such that $\hat{\tau}_0|_{\mathrm{Cl}(L(C))} \in \mathcal{A}$. Let us denote $\rho_0 := \hat{\tau}_0|_{\mathrm{Cl}(L(C_0))}$. Then $|\rho_0| \le |\mathrm{Cl}(L(C_0))| \le \mathrm{rk}(\neg C_0) \le r$. By the conditions of the theorem, there is a linearly described strategy of Delayer in the $(\varphi, \mathcal{A})$-game with starting position ρ_0 that guarantees him to earn $t - |\rho_0|$ white coins while paying at most c black coins.

Let Σ be the set of all assignments π such that π satisfies $\neg C_0$ and $\hat{\pi}|_{\mathrm{Cl}(L(C_0))} = \rho_0$. Since $\tau_0 \in \Sigma$, $\Sigma \neq \emptyset$.

Consider a random assignment $\sigma \in \Sigma$ and make $t - r$ steps in the linear branching program from C_0 according to σ. Notice that $t - r \le (t - |\rho_0|) + |\rho_0| - \mathrm{rk}(\neg C_0)$. Let C be the linear clause at the end of the path. By Theorem 4.3, with probability at least $2^{-(\ell-1)c}$, C is $\mathcal{A}$-good. By Lemma 2.19, C is not a clause of $\phi \circ g$, hence, $C \in S_{t-r}(C_0)$. Thus, $\mathrm{rk}(\neg C) \ge r(\ell+1) + c\ell$.

Let τ be a random full assignment of variables $\varphi \circ g$.

Let us estimate $\Pr[\tau \in \Sigma] \ge \Pr[\tau \text{ satisfies } \neg C_0, \forall i \in \mathrm{Cl}(L(C_0)), j \in [\ell], \tau(x_{i,j}) = \tau_0(x_{i,j})] \ge 2^{-r(\ell+1)}$. In the last inequality, we used that the event is defined by a satisfiable linear condition on τ of rank at most $\mathrm{rk}(\neg C_0) + \ell |\mathrm{Cl}(L(C_0))| \le r(\ell+1)$. Then

$$\Pr[\Pi \text{ contains a linear clause } C \text{ with } \mathrm{rk}(\neg C) \ge r + c\ell \text{ such that } \tau \text{ satisfies } \neg C] \ge$$
$$\Pr[\Pi \text{ contains a linear clause } C \text{ with } \mathrm{rk}(\neg C) \ge r + c\ell \text{ such that } \tau \text{ satisfies } \neg C \mid \tau \in \Sigma] \cdot$$
$$\Pr[\tau \in \Sigma] \ge 2^{-(\ell-1)c} 2^{-r(\ell+1)}.$$

If Ψ is a satisfiable linear system such that $\mathrm{rk}(\Psi) \ge r(\ell+1) + c\ell$, then $\Pr[\sigma \text{ satisfies } \Psi] \le 2^{-r(\ell+1)-c\ell}$. Hence, the refutation contains at least 2^c clauses C such that $\mathrm{rk}(\neg C) \ge r(\ell+1)+c\ell$. □

Let D_0 denote the empty clause from Π. If for every $\mathcal{A}$-good clause C such that there is a path from D_0 to C of length t, $\mathrm{rk}(\neg C)$ is at least $c(\ell+2)$, then (since $c(\ell+2) > c\ell$) by Claim 8.2, the size of the refutation Π is at least 2^c. Otherwise, there is an $\mathcal{A}$-good clause D_1 such that there is a path from D_0 to D_1 of length t and $\mathrm{rk}(\neg D_1) \leq c(\ell+2)$. Let $k := \lceil \log_{\ell+2}(\frac{t}{2c}) \rceil$, then $c(\ell+2)^{k-1} \leq t/2$. We repeat the same reasoning $k-1$ more times for all i from 1 to $k-1$ maintaining invariant $\mathrm{rk}(\neg D_i) \leq c(\ell+2)^i$: if for every $\mathcal{A}$-good clause C such that there is a path from D_i to C of length $t - c(\ell+2)^i$, $\mathrm{rk}(\neg C)$ is at least $c(\ell+2)^{i+1}$, then (since $c(\ell+2)^{i+1} > c(\ell+2)^i(\ell+1) + c\ell$) by Claim 8.2, the size of Π is at least 2^c. Otherwise, there is an $\mathcal{A}$-good clause D_{i+1} such that there is a path from D_i to D_{i+1} of length $t - c(\ell+2)^i$ and $\mathrm{rk}(\neg D_{i+1}) \leq c(\ell+2)^{i+1}$.

So we get that either the size of refutation is at least 2^c or depth is at least the length of the path from D_0 to D_1, from D_1 to D_2, etc, from D_{k-1} to D_k which is at least $kt/2 \geq \frac{t}{2}\log_{\ell+2}\left(\frac{t}{2c}\right)$. □

Corollary 8.3. *Let* $\mathrm{T}(G,c)$ *be an unsatisfiable Tseitin formula based on a spectral* (n,d,α)*-expander. Then, any* $\mathrm{Res}(\oplus)$ *refutation of* $\mathrm{T}(G,c) \circ Maj_5$ *has either size at least* 2^n *or depth at least* $n \cdot \left\lfloor d \cdot \frac{(1-\alpha)}{16} \right\rfloor \log_7 \lfloor d \cdot \frac{1-\alpha}{16} \rfloor$*. In particular, if* $d = \Theta(\log n)$ *and* $\alpha < 1$ *is a constant, then* $\mathrm{T}(G,c) \circ Maj_5$ *is a formula with* $m = 5dn/2$ *variables and of size* $\mathrm{poly}(m)$*. And any* $\mathrm{Res}(\oplus)$ *refutation of* $\mathrm{T}(G,c) \circ Maj_5$ *has either size at least* $2^{\Omega(m/\log m)}$ *or depth at least* $\Omega(m \log\log m)$*.*

Proof. Consider the set of partial assignments $\mathcal{T}$ and the natural strategy of Delayer in the $(\mathrm{T}(G,c),\mathcal{T})$-game defined in Subsection 4.2. Let us denote $t := \lfloor dn(1-\alpha)/8 \rfloor$. Consider some $\rho \in \mathcal{T}$, then the natural strategy of Delayer in the $(\mathrm{T}(G,c),\mathcal{T})$-game with starting position ρ guarantees him to earn at least $t - |\rho|$ white coins. The number of paid black coins does not exceed the possible number of connected components, which is at most n. By Lemma 4.2, this strategy is linearly described. The gadget Maj_5 is 2-stifling, hence, by Theorem 8.1 we get that size of any $\mathrm{Res}(\oplus)$ refutation of $\mathrm{T}(G,c) \circ Maj_5$ has either size at least 2^n or depth at least $\frac{t}{2}\log_7(t/2n) \geq n \cdot \left\lfloor d \cdot \frac{(1-\alpha)}{16} \right\rfloor \log_7 \lfloor d \cdot \frac{1-\alpha}{16} \rfloor$. □

9 Further Research

Chattopadhyay and Dvořák [8] recently established the following lifting theorem: if every resolution refutation of a CNF formula φ of width at most w requires depth at least d, then for every *strongly stifling* gadget g, any tree-like $\mathrm{Res}(\oplus)$ refutation of the lifted formula $\varphi \circ g$ of width at most w must have size at least 2^d. The class of strongly stifling gadgets forms a strict subclass of the class of 1-stifling gadgets.

Efremenko and Itsykson [14] recently extended this result to encompass all 1-stifling gadgets. Their approach builds on our game-based framework but introduces a refined tool—namely, the notion of *amortized closure*. This concept ensures that when a new linear form is added to a set, the closure can grow by at most one element. This refinement enables a more precise analysis in our proof of the depth-vs-size tradeoff, allowing the achievable depth to be improved up to $cn\log n$.

We consider it interesting to address the following questions connected with our research.

1. The lifting theorem from [9] applies to parity decision trees in the Boolean case. Developing a similar technique for strongly read-once branching programs [20] would be an interesting direction.

2. Lifting was used to establish the NP-hardness of automating algebraic proof systems [12]. Can the lifting approach from the paper similarly prove that regular Res($\oplus$) is NP-hard to automate?

3. Our lower bound for bounded-depth Res($\oplus$) applies only to lifted Tseitin formulas. Extending it to a broader class of formulas could be valuable, potentially aiding in separating unrestricted Res($\oplus$) from bounded-depth Res($\oplus$).

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