On the Partial Derivative Method Applied to Lopsided Set-Multilinear Polynomials

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Abstract

We make progress on understanding a lower bound technique that was recently used by the authors to prove the first superpolynomial constant-depth circuit lower bounds against algebraic circuits.

More specifically, our previous work applied the well-known partial derivative method in a new setting, that of *lopsided* set-multilinear polynomials. A set-multilinear polynomial $P \in \mathbb{F}[X_1, \ldots, X_d]$ (for disjoint sets of variables $X_1, \ldots, X_d$) is a linear combination of monomials, each of which contains one variable from $X_1, \ldots, X_d$. A lopsided space of set-multilinear polynomials is one where the sets $X_1, \ldots, X_d$ are allowed to have different sizes (we use the adjective "lopsided" to stress this feature). By choosing a suitable lopsided space of polynomials, and using a suitable version of the partial-derivative method for proving lower bounds, we were able to prove constant-depth superpolynomial set-multilinear formula lower bounds even for very low-degree polynomials (as long as d is a growing function of the number of variables N). This in turn implied lower bounds against general formulas of constant-depth.

A priori, there is nothing stopping these techniques from giving us lower bounds against algebraic formulas of *any* depth. We investigate the extent to which this lower bound can extend to greater depths. We prove the following results.

1. We observe that our choice of the lopsided space and the kind of partial-derivative method used can be modeled as the choice of a multiset $W \subseteq [-1, 1]$ of size d. Our first result completely characterizes, for any product-depth Δ, the best lower bound we can prove for set-multilinear formulas of product-depth Δ in terms of some combinatorial properties of W, that we call the *depth-Δ tree bias* of W.
2. We show that the maximum depth-3 tree bias, over multisets W of size d, is $\Theta(d^{1/4})$. This shows a stronger formula lower bound of $N^{\Omega(d^{1/4})}$ for set-multilinear formulas of product-depth 3, and also puts a non-trivial constraint on the best lower bounds we can hope to prove at this depth in this framework (a priori, we could have hoped to prove a lower bound of $N^{\Omega(\Delta d^{1/\Delta})}$ at product-depth Δ).
3. Finally, we show that for small Δ, our proof technique cannot hope to prove lower bounds of the form $N^{\Omega(d^{1/\operatorname{poly}(\Delta)})}$. This seems to strongly hint that new ideas will be required to prove lower bounds for formulas of unbounded depth.

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1 Introduction and Motivation

Basic background

This paper is motivated by questions arising in the area of *Algebraic Circuit complexity,* which studies the computational complexity of problems defined by families of multivariate polynomials. Given an infinite family of polynomials $(P_N(x_1, \ldots, x_N))_{N \geqslant 1}$ over a field $\mathbb{F}$, we consider the computational problem of evaluating P_N at input point $a \in \mathbb{F}^N$. Many natural and important computational problems can be stated in this language, including the problems of computing the determinant and the permanent, and that of multiplying matrices.

Algebraic circuits are succinct representations of multivariate polynomials that allow us to solve computational problems of the above form. More precisely, an algebraic circuit is a directed acyclic graph, where the sources are labelled by variables $x_1, \ldots, x_N$ or field elements and internal nodes (or *gates*) by algebraic operations $+$ and $\times$. Each internal node thus represents a polynomial in the variables $x_1, \ldots, x_N$ and a designated output gate represents the polynomial computed by the algebraic circuit. The *size* of the algebraic circuit is given by the number of gates. The *depth* and *product-depth* of an algebraic circuit denote the maximum number of gates and $\times$-gates respectively, on a directed path in the circuit.[1] Finally, we call an algebraic circuit an *Algebraic formula* if the underlying directed graph is a tree. (Equivalently, an Algebraic formula is just a nested algebraic expression made up of additions and multiplications, as one might write down on paper, represented in the form of a tree.)

An algebraic circuit for a polynomial P allows us to evaluate the polynomial P on a given input in time polynomially related to the size of the circuit. Thus, algebraic circuits are a restricted, but natural, model of computation for computational problems of this form. The study of this model of computation is one of the principal topics of study in Algebraic circuit complexity, and has received much attention over the past four decades (see e.g. [3, 26, 24] for nice introductions). Many central questions in Boolean circuit complexity have analogous and closely-related versions in the algebraic setting. For instance, the VP vs. VNP question [28], which is the problem of proving explicit lower bounds against algebraic circuits, is formally easier than the (non-uniform) P vs. NP question [2]. The problem of proving lower bounds against algebraic formulas is similarly closely related to the problem of proving lower bounds against the Boolean complexity class NC^1.

A recent result [18]

While circuit lower bounds in the algebraic setting are formally easier than the Boolean setting, they still have been hard to come by. For example, a famous line of research in the 1980s [1, 7, 13, 23, 27] showed exponential lower bounds against Boolean circuits of constant-depth, but did not yield such results for algebraic circuits.[2] This situation was somewhat rectified recently by the authors [18], building on some important earlier results in the area [20, 21]. In particular, we were able to prove superpolynomial lower bounds against constant-depth algebraic circuits over fields of characteristic zero.

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[1] W.l.o.g., we may assume that the product-depth and depth of a circuit are related to each other by a multiplicative factor of 2. However, some results are easier to state in terms of product-depth.

[2] Note that algebraic circuit lower bounds are not necessarily easier than Boolean circuit lower bounds in the constant-depth setting. However, some of these ideas did translate in the setting of constant-sized fields. [8, 9]

This paper is motivated by the problem of extending this lower bound to stronger models of computation. At a high level, our results are as follows.

- We show that our previous result [18] can be formulated purely in terms of a combinatorial property of the space of polynomials under consideration.
- We characterize the best lower bound that can be achieved in this framework at product-depth 3. It is better than the analogous lower bound from [18], but not as good as one might hope at first sight (as explained below).
- We place limitations on how well the bound extends to higher depths.

To describe these results in detail, we first recall the outline of the proof of [18].

The proof of [18]

The proof of [18] proceeds in two steps. In the first step, we reduce the problem of proving lower bounds for general circuits of depth Δ to proving lower bounds for *product-depth-*$(\Delta-1)$ circuits that have a special structure. In the second step, we prove lower bounds for the structured circuits. We describe these steps in some more detail next.

Step 1: Set Multilinearization. We work throughout with a partition of the variable set $X = \{x_1, \ldots, x_N\}$ into $X_1 \cup X_2 \cup \cdots \cup X_d$. Given such a partition, a *set-multilinear* monomial w.r.t. this variable partition is a monomial of degree d that contains exactly one variable from each of $X_1, X_2, \ldots, X_d$. A set-multilinear polynomial P is a linear combination of set-multilinear monomials. We denote the space of set-multilinear polynomials w.r.t. $X_1, \ldots, X_d$ by $\mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$. A set-multilinear circuit or formula is one where each gate computes a set-multilinear polynomial w.r.t. a subset of $\{X_1, \ldots, X_d\}$. An important example of a set-multilinear polynomial is the *Iterated Matrix Multiplication* polynomial $\mathrm{IMM}_{n,d}$, where $X_1, \ldots, X_d$ are square matrices of dimension $n \times n$ with distinct indeterminates, and the polynomial represents, say, the $(1,1)$th entry of the product of these matrices.

In the first step of the proof, we show that if a polynomial $P \in \mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$ has a circuit C of depth Δ and size s, then it also has a set-multilinear circuit C' of product-depth $\Delta - 1$ and size $s' = \mathrm{poly}(s) \cdot d^{O(d)}$. Note that while the blow-up in size in going from C to C' is large as a function of d, it can be made small (say $\mathrm{poly}(N)$) assuming that d is a slow-growing function of N (say, $d = O(\log N/\log\log N)$). So, to prove superpolynomial constant-depth circuit lower bounds, it suffices to prove superpolynomial lower bounds for constant-depth set-multilinear circuits in this *low-degree setting.*

Step 2: Set-multilinear lower bounds for low-degree polynomials. Lower bounds for constant-depth set-multilinear circuits have been known since the work of Nisan and Wigderson [20] from the 1990s. However, such lower bounds were typically of the form $\exp(d^{\Omega(1)}) \cdot \mathrm{poly}(N)$, which are not good enough for our purposes in the low-degree setting. The main contribution of [18] was to prove a lower bound of the form $N^{\omega_d(1)}$, which yields a superpolynomial lower bound for any degree $d = d(N)$ which is a growing function of N.

Somewhat surprisingly, the proof of this latter lower bound used just the lower bound technique of Nisan and Wigderson [20], which goes by the name of the *partial derivative method.* The key observation was to apply this technique to a suitable space of set-multilinear polynomials. Specifically, it is crucial in the proof to allow for the sets $X_1, \ldots, X_d$ to have fairly different sizes. To stress this feature, we refer to such a space of set-multilinear polynomials as *lopsided.*

For such polynomials that have efficient small-depth set-multilinear formulas, we argue that certain matrices associated to these polynomials have low rank. This is the basic recipe suggested by the partial derivative method, and is described in more detail later.

To complete the argument, we need to find explicit polynomials for which the associated matrices have high (ideally maximal) rank. We do this by considering suitable restrictions of $\mathrm{IMM}_{n,d}$ where $n = \max_{i\in[d]} |X_i|$. Using this idea, we showed [18] a lower bound of $N^{d^{\exp(-O(\Delta))}}$ for set-multilinear circuits of product-depth Δ. In conjunction with Step 1, this implies a superpolynomial lower bound for constant-depth algebraic circuits, and in fact for circuits of depth $o(\log\log d)$.

The potential of this lower bound technique

Can the above proof strategy be used to prove lower bounds for stronger models of computation, such as algebraic formulas of unbounded depth or, optimistically, even algebraic circuits? It turns out that Step 1 of the strategy still works, as shown in previous work of Nisan and Wigderson [20] and Raz [22]. Consequently, proving superpolynomial set-multilinear lower bounds against these models in the low-degree setting imply general circuit or formula lower bounds.

However, a problem arises because of the technique used in Step 2. As $\mathrm{IMM}_{n,d}$ (or more precisely, its restrictions) is a polynomial of "maximal complexity" for the partial derivative method, we cannot use it to prove lower bounds for computational models that can compute this polynomial efficiently. In particular, this suggests a new idea is required to prove lower bounds for, say, set-multilinear circuits of depth $O(\log d)$, which can compute $\mathrm{IMM}_{n,d}$ efficiently.

Nevertheless, this does not seem to rule out lower bounds for circuits of depth $o(\log d)$, or for formulas (of any depth). A simple, folklore divide-and-conquer strategy shows that $\mathrm{IMM}_{n,d}$ has set-multilinear circuits of product-depth Δ and size $n^{O(d^{1/\Delta})}$, and also set-multilinear formulas of product-depth Δ and size $n^{O(\Delta d^{1/\Delta})}$. Given the fact that this basic bound has not been improved upon significantly[3] for a long time, it is tempting to conjecture that it is tight, at least in the set-multilinear setting. If so, it seems that we could hope to prove lower bounds for set-multilinear circuits of depth $o(\log d)$ and formulas of any depth. Doing this would yield at least lower bounds for general algebraic formulas, which would be a very interesting result. This brings us to our main motivating question.

▶ **Question 1.** *Can we hope to use the partial derivative method (as applied to lopsided spaces of set-multilinear polynomials) to prove set-multilinear lower bounds that match the standard divide and conquer algorithms for* $\mathrm{IMM}_{n,d}$*?*

Our results in this paper indicate that the answer to this question is probably "No", and that, alone, the proof technique from [18] is not powerfull enough to handle formulas of depth $(\log d)^{o(1)}$. In the process of proving these results, we also introduce what we believe is a clean framework for studying the power of this technique.

We start with a more formal description of the partial derivative method and then state our results.

[3] A famous result of Gupta, Kamath, Kayal and Saptharishi [12] does improve this bound, but gives up on set-multilinearity. Moreover, the basic form of the bound is still preserved. More precisely, their work implies circuits of product-depth Δ and size $n^{O(d^{1/2\Delta})}$.

The partial derivative method for lopsided set-multilinear polynomials

We prove lower bounds for set-multilinear polynomials $P(X_1, \ldots, X_d)$ where each $|X_i| = n^{\alpha_i}$ for some $\alpha_i \in (0, 1]$. Given such a polynomial P, we associate with it a matrix as follows. We partition $[d]$ into sets $\mathcal{P}$ and $\mathcal{N}$. The rows of the matrix are associated with set-multilinear monomials over the variable sets $\{X_i \ : \ i \in \mathcal{P}\}$, and the columns symmetrically with the set-multilinear monomials over $\{X_j \ : \ j \in \mathcal{N}\}$. Given a row labelled by monomial m_1 and a column labelled by monomial m_2, the corresponding entry in the matrix is the coefficient of the set-multilinear monomial $m_1 m_2$ in the polynomial P. We use the rank of this matrix (or, more precisely, how close it is to full-rank) to prove lower bounds on the algebraic circuit complexity of P.

We define this more precisely now. Note that the matrix is completely specified by the choice of the numbers $\alpha_1, \ldots, \alpha_d$ and the partition $[d] = \mathcal{P} \cup \mathcal{N}$. We can describe these together by the multiset $W \subseteq [-1, 1]$, defined by $W = \{\alpha_i \ : \ i \in \mathcal{P}\} \cup \{-\alpha_j \ : \ j \in \mathcal{N}\}$. Finally, we use $M_W(P)$ to denote the above matrix.

Note that $M_W(P)$ is a matrix with $R = n^{\sum_{\alpha \in W \cap (0,1]} \alpha}$ rows and $C = n^{\sum_{\alpha \in W \cap [-1,0)} |\alpha|}$ columns. In particular, the rank of the matrix $M_W(P)$ is bounded by the minimum of these quantities. We consider the *relative rank* of P, defined as follows.

$$\mathrm{relrk}_W(P) = \frac{\mathrm{rank}(M_W(P))}{\sqrt{RC}} = \frac{\mathrm{rank}(M_W(P))}{n^{\frac{1}{2}\sum_{\alpha \in W} |\alpha|}}. \tag{1}$$

Observe that the quantity in the denominator is the geometric mean of the number of rows and the number of columns of $M_W(P)$ and hence $\mathrm{relrk}_W(P) \in [0, 1]$. In fact, more generally, it is not hard to see that as $\mathrm{rank}(M_W(P)) \leqslant \min\{R, C\}$, we have $\mathrm{relrk}_W(P) \leqslant n^{-|\sum_{\alpha \in W} \alpha|/2}$.

Further, it was shown by the authors [18] that for any W, there is a polynomial P_0 such that $\mathrm{relrk}_W(P_0) = n^{-|\sum_{\alpha \in W} \alpha|/2}$ and P_0 can be obtained by starting with an instance of $\mathrm{IMM}_{\mathrm{poly}(n),d}$ and setting some variables to 0 and identifying variables within certain sub-matrices, i.e. by a *set-multilinear projection.*

High-level description of the results

Our results give a better understanding of what lower bounds the partial derivative method can hope to show in this setting.

- Our first main result is a transformation of our problem to a combinatorial problem about labelled trees. More precisely, we show that understanding the best lower bound our techniques can hope to prove in the low-degree setting is perfectly captured by the best-possible "tree-like decomposition" of the set W.[4]
  While this transformation is simple, it is conceptually clean, and simplifies the problem in multiple ways. Firstly, it eliminates the parameter n (which is roughly the number of variables in the underlying polynomial) and makes completely clear the dependence of the lower bound on properties of the multiset W. Secondly, this reformulation of the problem completely eliminates any mention of polynomials or algebra from the problem. It is now purely a problem about the "additive structure" of W.
- Our second result uses the above characterization of the problem to give a near-perfect understanding of the best lower bounds we can prove for set-multilinear formulas of product-depth 3 (i.e. $\Sigma\Pi\Sigma\Pi\Sigma\Pi\Sigma$ formulas). More precisely, we show that the best product-depth-3 lower bound we can prove via our proof technique is $n^{\Theta(d^{1/4})}$. This is interesting for the following two different reasons.

[4] This is not to be confused with standard tree decompositions of graphs, which have no connection with objects studied here.

For one, this is a stronger lower bound than known previously for set-multilinear formulas of product-depth 3 in the low-degree regime: Nisan and Wigderson [20] showed a lower bound of $\exp(\Omega(d^{1/3})) \cdot \mathrm{poly}(N)$ (which does not yield anything for $d = O(\log N)$), while in our earlier work [18], we showed lower bounds of $n^{\Omega(d^{1/7})}$.
On the other hand, the result also implies that this technique does not go as far as we would like. Recall from above that the (suspected) optimal lower bound for $\mathrm{IMM}_{n,d}$ at product-depth 3 is $n^{\Omega(d^{1/3})}$. So, our result implies that this technique cannot be used to obtain this bound at product-depth 3.

- The above results already indicate that we are not able to prove the best possible lower bound we could hope for product-depth-3 set-multilinear formulas. However, it is still conceivable that we can hope to prove a lower bound which stays "close" to the right expected bound for $\mathrm{IMM}_{n,d}$ (say a bound of the form $n^{\Delta d^{\Omega(1/\Delta)}}$), which could as yet lead to superpolynomial formula lower bounds.
In our third result, we give strong indication that this is not the case, by showing that this technique cannot prove lower bounds of the form $n^{d^{1/\Gamma(\Delta)}}$ for a quasipolynomial function $\Gamma(\cdot)$, and small enough Δ.

1.1 Formal description of the results

To describe the results formally, we introduce a combinatorial measure of the complexity of the multiset $W \subseteq [-1, 1]$. In the low-degree setting, this will characterize the best lower bound we can prove via our lower bound technique.

Notation

Let $W \subseteq \mathbb{R}$ be a multiset. Throughout $|W|$ denotes the size of the multiset (i.e. counted with multiplicity) and $\mathrm{Sum}(W)$ denote the sum of its elements. Finally, $\|W\|_1$ denotes the L_1-norm of W (i.e. the sum of the absolute values of the elements of W).

▶ **Definition 2** (W-trees, path bias, tree bias)**.** *Let $W = \{\alpha_1, \dots, \alpha_d\}$ be a multiset contained in $[-1, 1]$. A W-*tree *T, or equivalently* a tree T for W, *is a rooted, directed tree*[5] *with $d = |W|$ leaves which are labelled by distinct elements of the form (i, α_i) ($i \in [d]$).*[6] *Any vertex v of T thus corresponds to a subset W_v of W (corresponding to the leaves of the subtree induced by v) and we define $\mathrm{Sum}(v)$ to be $\mathrm{Sum}(W_v)$.*

An internal path *π in T is a path from the root to an internal (i.e. non-leaf) node. Given such an internal path π, we define the set of* Off-path nodes *of π, denoted $\mathrm{Offpath}(\pi)$ to be the set of nodes v of the tree T that are not on the path π, but have a parent on the path π. We define* the bias of the path π, *denoted $\mathrm{bias}(\pi) = \left(\sum_{v \in \mathrm{Offpath}(\pi)} |\mathrm{Sum}(v)|\right) - |\mathrm{Sum}(r)|$ where r is the root of T.*

(It is easy to check that if π is any internal path, then $W = W_r$ is the disjoint union of W_v ($v \in \mathrm{Offpath}(\pi)$). Hence, by the triangle inequality, we have $|\mathrm{Sum}(r)| \leqslant \sum_{v \in \mathrm{Offpath}(\pi)} |\mathrm{Sum}(v)|$. Thus, $\mathrm{bias}(\pi) \geqslant 0$ for any internal path π.)

Finally, we define the path bias of T w.r.t. W, *denoted $\mathrm{Pathbias}_W(T)$, to be the* maximum *bias of any internal path of T. If the tree T has depth 0 (i.e. it consists of just the root node), then we define the path bias of T w.r.t. W to be 0.*

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[5] The edges are directed away from the root.

[6] We require the label to be a pair here as W is a multiset where elements may repeat. If the elements of W are all distinct, then we can think of the labels as simply elements of W.

With the above notation in place, we can define the combinatorial measure mentioned above. We define the depth-Δ tree bias of W *to be the* minimum *path bias of any depth-Δ W-tree T. We denote this quantity by* $\text{Treebias}_\Delta(W)$.

Our first theorem relates the depth-Δ tree bias of $W = \{\alpha_1, \ldots, \alpha_d\} \subseteq [-1, 1]$ with the best lower bound we can prove using the complexity measure $\text{relrk}_W(\cdot)$.

▶ **Theorem 3** (Connecting tree bias with relative rank). *Let n, d be positive integer parameters.*[7] *Let $\Delta \geqslant 1$ be any integer. Assume $W \subseteq [-1, 1]$ is a multiset of size d such that* $\text{Treebias}_\Delta(W) = t$. *Then, for any set-multilinear formula F of product-depth at most Δ and size at most s, we have* $\text{relrk}_W(F) \leqslant (d^{3d} \cdot s \cdot n^{-t/2}) \cdot n^{-|\text{Sum}(W)|/2}$. *Conversely, for any n and d, there is a set-multilinear formula F with at most $3^d n^{t/2}$ leaves and of product-depth Δ such that* $\text{relrk}_W(F) \geqslant 2^{-d} \cdot n^{-|\text{Sum}(W)|/2}$.

This theorem is the consequence of Lemmas 13 and 14 and will be proved in Section 3. As already noted above, for any polynomial $P \in \mathbb{F}_{sm}[X_1, \ldots, X_d]$ (with $|X_i| = n^{|\alpha_i|}$ for each $i \in [d]$), we have $\text{relrk}_W(P) \leqslant n^{-|\text{Sum}(W)|/2}$. Theorem 3 shows that this maximum possible relative rank can be achieved by product-depth-Δ formulas of size $n^{O(t)}$, but not those of size $n^{o(t)}$, where $t = \text{Treebias}_\Delta(T)$. This means that the best lower bound we can hope to prove via this technique is $n^{\Theta(t)}$.

The next couple of theorems give an understanding of the maximum possible tree bias for various depths Δ. The first result gives tight bounds on the maximum possible tree bias of a given multiset W for depth 3 (Section 4 will be dedicated to this result).

▶ **Theorem 4** (Tight bounds on tree bias for depth 3). *Let d be a growing integer parameter. Then,* $\max_W \text{Treebias}_3(W) = \Theta(d^{1/4})$, *where W ranges over multisets from $[-1, 1]$ of size d in the expression above.*

The second result (proved in the long version of the paper) gives an asymptotic bound for larger depths (as long as Δ is bounded by a small function of d).

▶ **Theorem 5** (Bounds on tree bias for larger depths). *Let d, Δ be growing integer parameters with $\Delta = 2^{o(\sqrt{\log \log d})}$. Then, we have* $\max_W \text{Treebias}_\Delta(W) \leqslant d^{1/\Delta^{\Omega(\log \Delta)}}$, *where W ranges over multisets from $[-1, 1]$ of size d.*

1.2 Proof Outline

Throughout this section, we work with a multiset $W = \{\alpha_1, \ldots, \alpha_d\} \subseteq [-1, 1]$ and a space of lopsided set-multilinear polynomials $\mathbb{F}_{\text{sm}}[X_1, \ldots, X_d]$ where $|X_i| = n^{|\alpha_i|}$. Recall also that we are working in the low-degree setting, i.e. d is a slow-growing function of n. All formulas in this section should be assumed to be set-multilinear.

Motivation for tree bias

We start by motivating the notion of tree bias which, at first sight, might appear mysterious to the reader. In fact, this notion comes up quite naturally in the course of constructing small set-multilinear formulas that have large relative rank. These constructions, in turn, are motivated by the following basic properties of relative rank which are all slight modifications of standard facts used in the literature. In this form they can be found in our earlier work [19].[8]

[7] We think of d as a slow-growing function of n.

[8] The paper deals with a related notion of relative rank w.r.t. *ordered* W (or equivalently, W is replaced by a tuple $(\alpha_1, \ldots, \alpha_d)$). However, the proof works in the same way for multisets.

▶ **Lemma 6** (Properties of Relative Rank).

1. *(Imbalance) Say* $P \in \mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$*. Then,* $\mathrm{relrk}_W(P) \leqslant n^{-|\mathrm{Sum}(W)|/2}$.
2. *(Sub-additivity) Say* $P, Q \in \mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$*. Then* $\mathrm{relrk}_W(P+Q) \leqslant \mathrm{relrk}_W(P) + \mathrm{relrk}_W(Q)$.
3. *(Multiplicativity) Say* $P = P_1 \cdot P_2 \cdot \ldots \cdot P_t$ *and assume that for each* $i \in [t]$, $P_i \in \mathbb{F}_{\mathrm{sm}}[X_j : j \in S_i]$*, where* $\{S_1, \ldots, S_t\}$ *is a partition of* $[d]$*. Then* $\mathrm{relrk}_W(P) = \mathrm{relrk}_W(P_1 \cdot P_2 \cdot \ldots \cdot P_t) = \prod_{i \in [t]} \mathrm{relrk}_{W_i}(P_i)$*, where* $W_i = \{\alpha_j \mid j \in S_i\}$.

With these properties in mind, we try to construct small set-multilinear formulas with optimally large relative rank. We do not lose much generality in assuming that $\mathrm{Sum}(W) \approx 0$, which we will do in the rest of this proof outline. So, the optimal relative rank is 1.

It is instructive to consider the example of W such that $\alpha_1 = \cdots = \alpha_{d/2} = 1$ and $\alpha_{d/2+1} = \cdots = \alpha_d = -1$. We start with a trivial formula F that consists of a single variable $x_1 \in X_1$, which has relative rank $n^{-1/2}$. Does it make sense to take linear combinations of such formulas? From the perspective of relative rank, the answer is No, because that increases the size without increasing the relative rank at all, by the Imbalance criterion in Lemma 6. So we can only multiply variables (from different sets, as we are dealing with set-multilinear formulas). Moreover, it makes sense to multiply variables such that the corresponding α_is have different signs, as multiplying variables from X_1 and X_2 (say) would only make the imbalance worse. So we multiply $x_1 \in X_1$ and $x_{d/2+1} \in X_{d/2+1}$. This creates a formula of relative rank $1/n$, by the property of Multiplicativity. By Sub-additivity, we need to sum at least n such formulas to get a formula of relative rank 1 (which is optimal). And indeed, this can be done, say, with an inner product between the variables of X_1 and $X_{d/2+1}$. Multiplying $d/2$ such formulas together (for a partition of $\alpha_1, \ldots, \alpha_d$ into positive and negative pairs) gives us a formula of product-depth 2, size $O_d(n)$, and relative rank 1.[9] One can see that the underlying multiplicative structure of the formula thus constructed naturally suggests a W-tree T of the form shown in Figure 1. This is a W-tree of depth-2 and bias 2 (which is the best possible for this W).

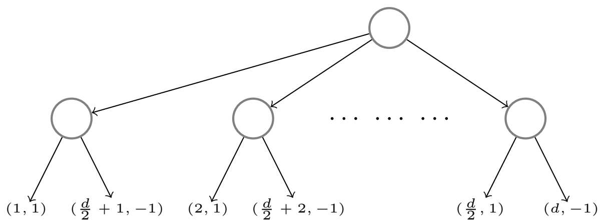

Figure 1 The W-tree of depth 2 and bias 2 arising from the formula construction above.

The above indicates a general technique for constructing formulas of large relative rank. Start by finding a $W' \subseteq W$ such that $|\mathrm{Sum}(W')|$ is small. Construct a formula of plausibly optimal relative rank (i.e. $n^{-|\mathrm{Sum}(W'))|/2}$) over the variable sets corresponding to W' by adding enough set-multilinear monomials so that sub-additivity no longer indicates that the rank of the formula is small. In doing this, we end up taking a sum of size n^b where

$$b := \frac{1}{2} \sum_{i \in W'} |\alpha_i| - \frac{|\mathrm{Sum}(W')|}{2}. \tag{2}$$

[9] This is an example of Nisan and Wigderson [20], aptly called the *Product of Inner Products* polynomial.

This indicates that it helps to take W' to be a small set, since otherwise this formula would be too large (if there were no such constraint, we could simply have taken $W' = W$). We partition W into small sets $W'_1, \ldots, W'_r$ this way, and construct formulas for each. Then, applying again the same principle to the multiset $\{\mathrm{Sum}(W'_1), \ldots, \mathrm{Sum}(W'_r)\}$, we get a high-rank set-multilinear formula over all of $X_1, \ldots, X_d$. As in the simple example above, this gives rise to a multiplicative structure that can be described by means of a W-tree T. The set W' constructed above corresponds to one of the nodes at height 1 in T and the quantity b in (2) is (almost) something we will define later to be the bias of the corresponding node, and $n^{b/2}$ lower bounds the size of the constructed formula. However, a careful analysis of the construction shows that the size of the formula is actually larger: at each node of the tree T, the formula uses a sum governed by the bias of the corresponding node. This naturally ends up yielding a formula whose size is governed by the path bias of T. Minimizing this over the choice of all trees yields the tree bias of W, as defined above.

Proof of Theorem 3

The above outline already indicates how to construct a set-multilinear formula of product-depth Δ and size $n^{O(\mathrm{Treebias}_\Delta(W))}$ that computes a polynomial of optimal relative rank. The only part that is unclear is how to ensure that the bounds on relative rank imposed by sub-additivity are actually tight. We do this by a careful inductive definition of the formulas. In a revision of our earlier paper [18], we showed how to do this for a specific W which contains only the two distinct elements -1 and $1/\sqrt{2}$. In this paper, we extend this construction to all W. This gives the second part of Theorem 3.

In the process, we note that the formulas we construct all have a special property: they have a *unique* multiplicative structure, i.e. they build up all their set-multilinear monomials in the same way, given by a single W-tree T. In principle, a general set-multilinear formula could contain many different kinds of trees (e.g. by summing formulas corresponding to different trees). These special formulas that we construct have been studied before: they are called *Pure* formulas [20] or *Unique Parse Tree* (UPT) formulas [16, 15]. We use the latter terminology.

For the first part of Theorem 3, we proceed as follows. We first show that UPT formulas of product-depth Δ have indeed the claimed upper bound on the relative rank, by using the basic properties of relative rank from Lemma 6 and a simple inductive argument. To argue about a general set-multilinear formula F, we show that any set-multilinear formula can be written as a sum of $O_d(1)$ many UPT formulas of the same size and product-depth. Using the sub-additivity of relative rank and the bound for UPT formulas, we see that F also has small relative rank.

We illustrate the power of the latter theorem with a short proof of one of the main results of [18]: an $n^{\Omega(\sqrt{d})}$ lower bound for set-multilinear formulas of product-depth 2.[10] By Theorem 3, it suffices to construct a multiset $W \subseteq [-1, 1]$ with $|\mathrm{Sum}(W)| = 0$ and tree bias $\Omega(\sqrt{d})$. Consider a W with $\Theta(d)$ copies each of (-1) and $\alpha := (1 - 1/\sqrt{d})$ so that $\mathrm{Sum}(W) = 0$. Given any depth-2 W-tree T, it can be checked that one of the following hold.

- There is a depth-1 vertex u with $t_u \geqslant \sqrt{d}/2$ children. In this case, any path through u has bias $\Omega(\sqrt{d})$.
- Every u at depth-1 has $t_u < \sqrt{d}/2$ children, in which case $|\mathrm{Sum}(u)| \geqslant t_u/(2\sqrt{d})$. This implies that any path in T has bias $\sum_u t_u/(2\sqrt{d}) = \sqrt{d}/2$.

[10] This is essentially the heart of the argument of [18], abstracting away the details about algebraic formulas, and keeping only the combinatorial core.

Proof of Theorem 4

In a similar way, we can also extend the results of [18] to show improved lower bounds for product-depth 3 (i.e. $\Sigma\Pi\Sigma\Pi\Sigma\Pi\Sigma$ formulas). More precisely, taking W as above but redefining $\alpha = 1-(1/d^{1/4})-(1/d^{3/4})$, we are able to prove a tree-bias lower bound of $\Omega(d^{1/4})$. This implies a formula lower bound of $n^{\Omega(d^{1/4})}$, which improves upon a lower bound of $n^{\Omega(d^{1/7})}$ from our previous work.

In the second part of the proof, we show that this is the best bound that this technique can prove, for *any* choice of W. Equivalently, we can show that every W has depth-3 tree bias $O(d^{1/4})$. We illustrate the idea with a sketch of the special case when W has two distinct elements (as in the two lower bounds above). In this case, it is not hard to argue that without loss of generality, the two distinct elements of W are (-1) and $\alpha \in (0,1]$.

First of all, we observe that any W has a tree of depth Δ and path bias $O(\Delta\|W\|_1^{1/\Delta})$, where $\|W\|_1$ denotes the sum of the absolute values of the elements of W. This is analogous to the fact that $\mathrm{IMM}_{n,d}$ has set-multilinear formulas of depth-Δ and size $n^{O(\Delta d^{1/\Delta})}$. Call this the "basic construction".

Now, given W as above, we proceed as follows. By a classical result of Dirichlet (see, e.g. [14, Theorem 4.9]), for any t, there exist integers $q \in [t]$ and $p \in \{0,\ldots,t\}$ such that $|q\alpha - p| \leqslant 1/t$. Note that this gives a multiset $W' \subseteq W$ of size $p+q$ such that $|\mathrm{Sum}(W')| \leqslant 1/t$. We apply this result with $t = \sqrt{d}$ and proceed in one of two ways depending on the value of $p+q$.

- If $p+q \geqslant d^{1/4}$, then we can partition W into at most $r \leqslant d^{3/4}$ sets $W_1,\ldots,W_r$ of size $p+q$, each of which has sum at most $1/\sqrt{d}$. As $p+q \leqslant 2\sqrt{d}$, using the basic construction of depth 2, we get a tree T_i of bias $O(d^{1/4})$ for each W_i. Attaching all these to a common root gives a tree of path-bias $O(d^{1/4})$ (the root adds at most $d^{3/4}\cdot(1/\sqrt{d}) = d^{1/4}$ to the bias of any path).
- If $p+q \leqslant d^{1/4}$, then by using $d^{1/4}/(p+q)$ many disjoint sets of sum $1/\sqrt{d}$ each, we get a set W' of size $d^{1/4}$ and sum at most $d^{1/4}/((p+q)\cdot\sqrt{d}) \leqslant 1/d^{1/4}$. We partition W into $r \leqslant d^{3/4}$ sets $W_1',\ldots,W_r'$ of this form. We use a tree T_i of depth-1 for each W_i' (which has path bias at most $d^{1/4}$ trivially) and attach these to the leaves of a depth-2 tree for the set $\tilde{W} = \{\mathrm{Sum}(W_1),\ldots,\mathrm{Sum}(W_r)\}$. The latter tree is constructed using the basic construction of depth 2, and has bias $O(d^{1/4})$ as $\|\tilde{W}\|_1 \leqslant r/d^{1/4} \leqslant \sqrt{d}$.

This gives the argument in the case of W with only two distinct elements. For general W, we use a similar high-level argument. However, we need a suitable replacement for Dirichlet's theorem, which only works for the special W dealt with above. We prove a generalization of this theorem (see Lemma 9 below) to the setting of arbitrary multisets W. We think the statement is natural and interesting in its own right, but could not find mention of it in the literature.

In the special case that W contains d copies of $\alpha \in (0,1)$ and d copies of -1, the above implies the standard Dirichlet theorem used above. With the above generalized theorem in place, we can follow the structure of the argument for the special case, with technical modifications. This yields the depth-3 relative rank upper bound for any W.

Proof of Theorem 5 for depth Δ

While the proof of this theorem employs the same high-level argument as Theorem 22 described above, it is considerably more technical. We illustrate the idea again with the case when W contains only two distinct elements, which we can assume to be -1 and some $\alpha \in [0,1]$. Let $\mathrm{Bias}(\Delta, d)$ denote the largest possible bias of a depth-Δ W-tree. We give a constructive bound on this quantity by an inductive construction (based on Δ).

For $\Delta = 1$, we have the trivial bound $\text{Bias}(1, d) \leqslant d$. For $\Delta > 1$, we use Dirichlet's theorem to find integers $p, q \leqslant d^{1-\varepsilon}$ such that $|q - p\alpha| \leqslant d^{-(1-\varepsilon)}$. This gives us a set $W' \subseteq W$ of size $p + q$ such that $|\text{Sum}W'| \leqslant d^{-(1-\varepsilon)}$. There are again two cases to consider based on the magnitude of q.

- If $q \geqslant d^{\varepsilon}$, then this yields that $|W'| \geqslant d^{\varepsilon}$. Partitioning W into $t = d^{1-\varepsilon}$ subsets $W'_1, \ldots, W'_t$ of this form and using a recursive construction for each of $W'_1, \ldots, W'_t$, we get a W-tree of bias $\text{Bias}(\Delta - 1, d^{\varepsilon}) + O(1)$. (Here, the last $O(1)$ term accounts for the bias accrued at the root, which is only a constant.)
- Conversely, if $q \leqslant d^{\varepsilon}$, we pick as many sets $W'_1, \ldots, W'_r$ as we can to form a set W'' of size (roughly) $d^{1-\varepsilon}$. Note that $|\text{Sum}(W'')| \leqslant d^{1-\varepsilon}/d^{1-\varepsilon} \leqslant 1$. We partition W into $s \leqslant d^{\varepsilon}$ sets $W''_1, \ldots, W''_s$ of this form. We can construct a W-tree T of depth $\Delta = \Delta_1 + \Delta_2$ by
  - Constructing a W''_i-tree T_i of depth Δ_1 by constructing W'_j-tree $T_{i,j}$ of depth $\Delta_1 - 1$ for each $W'_j \subseteq W''_i$ and connecting these trees to a common root.
  - Constructing a depth-Δ_2 $\tilde{W}$-tree $\tilde{T}$, where $\tilde{W} = \{\text{Sum}(W''_1), \ldots, \text{Sum}(W''_s)\}$[11] and replacing the leaf labelled i with the tree T_i.

  As the sets $\tilde{W}$ and W'_i have size d^{ε} each, it makes sense to take $\Delta_1 = \Delta_2 = \Delta/2$. This leads to a bound on the bias of the tree T of $2 \cdot \text{Bias}(\Delta/2, d^{\varepsilon}) + O(1)$.

We choose ε to balance the bias obtained from each of the above two strategies. It is clear that if $\varepsilon < 1/(2\Delta)$ (say), then the first strategy yields a bad bound of $d^{1/2}$ (or worse). This implies that we must take $\varepsilon \geqslant 1/2\Delta$, which can yield a best possible upper bound of $d^{1/\Delta^{O(\log \Delta)}}$ from the second strategy. We show that this upper bound is indeed achievable, by taking $\varepsilon = \Theta(\log^2 \Delta/\Delta)$.

1.3 Related Work

Barriers for lower bound techniques

The partial derivative method and its variants have been used to prove several lower bounds in algebraic complexity theory including the recent work of the authors. While these techniques have been quite useful, it is unclear whether they can be used to separate VP from VNP. In the last decade, there were many attempts at understanding the limitations of these lower bound techniques. This has led to a body of work about *barrier* results [25, 10, 6, 11, 5, 4] in algebraic complexity theory. These results typically consider a large family of lower bound techniques and argue that such techniques cannot be used to prove strong lower bounds. However, all such results are either conditional, or hold for relatively weak models of computation (such as set-multilinear formulas of product-depth 1). In contrast to these results, here we focus on a specific technique, namely the technique that gave the first super-polynomial lower bound for low-depth circuits. We show an unconditional limitation on this technique with respect to a reasonably strong model of computation. Hence, our work is incomparable to this literature.

[11] There is a small technical point here, which is that we will be left with a few more elements not covered by any of the W''_is. We ignore this here.

Our other recent work [17]

In a different recent paper, we prove algebraic formula lower bounds for formulas of larger depths. Specifically, we are able to prove superpolynomial set-multilinear formula lower bounds for $\mathrm{IMM}_{n,n}$ and *non-commutative* formula[12] lower bounds for formulas of depths up to $o(\sqrt{\log d})$. Note that the first of these results is a lower bound in the *high-degree* setting. This does not immediately imply a lower bound for general formulas, as we do not know of an efficient transformation to set-multilinear formulas when the degree is large. The second result does not imply any lower bounds in the commutative setting, as far as we know. The results of this paper are thus somewhat orthogonal, as they apply to set-multilinear (commutative) formulas in the low-degree setting.

Organization

We start with some preliminaries in Section 2. We then prove Thorems 3 and 4 in Sections 3 and 4 respectively. The proofs of Theorem 5 and many other statements are deferred to the full version for lack of space.

2 Basic Preliminaries and Results from Previous Work

Fix any multiset $W = \{\alpha_1, \ldots, \alpha_d\} \subseteq [-1, 1]$ and let $\mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$ be a lopsided set-multilinear space of polynomials with $|X_i| = n^{\alpha_i}$.

The following is a consequence of earlier work of the authors.

▶ **Lemma 7** (Lower bounds from relative rank, Implicit in [18]). *Let d and n be integer parameters. Assume that $W = \{\alpha_1, \ldots, \alpha_d\} \subseteq [-1, 1]$ is an arbitrary multiset and consider the space $\mathbb{F}_{sm}[X_1, \ldots, X_d]$ where $|X_i| = n^{|\alpha_i|}$. Assume that we have shown the following: for any set-multilinear formula F (over variable sets $X_1, \ldots, X_d$) of size at most $s(n, d)$ and product-depth at most Δ, we have*

$$\mathrm{relrk}_W(F) \leqslant C_d \cdot \varepsilon_n \cdot n^{-|\mathrm{Sum}(W)|/2},$$

where C_d depends only on d and $\varepsilon_n \to 0$ as $n \to \infty$.

Then, for n large enough in comparison to d, any set-multilinear formula F of product-depth Δ computing $\mathrm{IMM}_{\mathrm{poly}(n),d}$ must have size at least $s(n, d)$. Further, any (possibly non-set-multilinear) formula of depth at most $\Delta + 1$ computing $\mathrm{IMM}_{\mathrm{poly}(n),d}$ must have size at least $s(n, d)/d^{O(\Delta d)}$.

The following simple proposition regarding path bias will be useful.

▶ **Proposition 8.** *Let $W \subseteq [-1, 1]$ be any finite multiset and let T be a W-tree with internal vertex u. If u has children $u_1, \ldots, u_r$, then*

$$\mathrm{Pathbias}_{W_u}(T_u) = \left(\max_{i\in[r]} \mathrm{Pathbias}_{W_{u_i}}(T_{u_i})\right) + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)|$$

where T_v denotes the subtree rooted at vertex v (which is, by definition, a W_v-tree in the natural way).

[12] This means that the operations of the formula are those of the non-commutative polynomial ring $\mathbb{F}\langle x_1, \ldots, x_N\rangle$ where variables do not commute.

Proof. Let p_v denote $\mathrm{Pathbias}_{W_v}(T_v)$ for any vertex v of T.

For any $i \in [r]$, let π_{u_i} denote the path of bias p_{u_i} in T_{u_i}. Let π_u denote the path in T_u obtained by adding the vertex u to π_{u_i}. Note that the off-path nodes of π_u are precisely the off-path nodes of π_{u_i} along with u_j ($j \neq i$). Thus, the bias of π_u can be written as

$$\begin{aligned}\mathrm{bias}(\pi_u) &= \left(\sum_{v \in \mathrm{Offpath}(\pi_u)} |\mathrm{Sum}(v)|\right) - |\mathrm{Sum}(u)| \\ &= \mathrm{bias}(\pi_{u_i}) + |\mathrm{Sum}(u_i)| + \sum_{j \in [r]\setminus\{i\}} |\mathrm{Sum}(u_j)| - |\mathrm{Sum}(u)| \\ &= p_{u_i} + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)|.\end{aligned}$$

As this holds for each $i \in [r]$, we have shown that

$$p_u \geqslant \left(\max_{i \in [r]} p_{u_i}\right) + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)|. \tag{3}$$

For the reverse inequality, we proceed in the same way. Let π_u be a path in T_u of bias p_u. If π_u has length 0, then we have

$$p_u = \mathrm{bias}(\pi_u) = \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)| \leqslant \left(\max_{i \in [r]} p_{u_i}\right) + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)|$$

and hence we are trivially done. Otherwise, the path π_u passes through some child u_i of u. Let π_{u_i} be the path in T_{u_i} obtained by removing u from π_u. Then, through the same sequence of equalities proved above, we get

$$p_u = p_{u_i} + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)| \leqslant \left(\max_{i \in [r]} p_{u_i}\right) + \left(\sum_{j=1}^{r} |\mathrm{Sum}(u_j)|\right) - |\mathrm{Sum}(u)|.$$

Hence, we have proved the reverse inequality to (3) and we are done. ◀

2.1 A Generalized form of Dirichlet's theorem

Here we prove a generalized form of the standard Dirichlet Principle (see, e.g. [14, Theorem 4.9]), which we will use in Sections 4 and in the proof of Theorem 5 in the full version.

▶ **Lemma 9** (A Generalized Form of the Dirichlet Principle). *Assume $d \geqslant 2$. Let $W \subseteq [-1, 1]$ be any multiset with at least d non-negative and d non-positive elements. Then, for each positive integer $t \leqslant 2d$, there is a multiset $T \subseteq W$ of size at most t such that $|\mathrm{Sum}(T)| \leqslant 4/(t-1)$.*

Proof. The proof is via the Pigeonhole principle. Fix a t as above and let $\ell = \lfloor t/2 \rfloor$. If W contains an element x such that $|x| \leqslant 2/\ell$, then we are done trivially, so we assume that this is not the case.

Let $\{x_1, \ldots, x_\ell\}$ and $\{-y_1, \ldots, -y_\ell\}$ be any ℓ positive and negative elements of W respectively (here, $x_i, y_i \in (2/\ell, 1]$ for each i).

For $i \in \{0, \ldots, \ell\}$, define $u_i = \sum_{j=1}^{i} x_j$ and $v_i = \sum_{j=1}^{i} y_j$. For $i, j \in \{0, \ldots, \ell\}$, let $w_{i,j} = u_i + v_j$. Note that as $x_i, y_i \in [0, 1]$ for each $i \in [\ell]$, we have $u_i, v_j \in [0, \ell]$ and $w_{i,j} \in [0, 2\ell]$ for each $i, j \in \{0, \ldots, \ell\}$. Also note that $u_0, \ldots, u_\ell$ and $v_0, \ldots, v_\ell$ are increasing sequences in which the difference between any pair of elements is strictly more than $2/\ell$.

Divide the interval $[0, 2\ell]$ into ℓ^2 sub-intervals of length $2/\ell$ each. By the pigeonhole principle, there exist distinct (i, j) and (i', j') from $\{0, \ldots, \ell\} \times \{0, \ldots, \ell\}$ such that $w_{i,j}$ and $w_{i',j'}$ lie in the same interval. In particular, we have

$$|w_{i,j} - w_{i',j'}| = |(u_i - u_{i'}) - (v_{j'} - v_j)| \leqslant \frac{2}{\ell}. \tag{4}$$

Fix such (i, j) and (i', j'). Since these pairs are distinct, they must differ in some coordinate. We assume that they differ in the first coordinate (the other case is similar).

Without loss of generality, assume that $i > i'$. We note that it cannot be the case that $j \geqslant j'$. This is because we would then have

$$|w_{i,j} - w_{i',j'}| = (u_i + v_j) - (u_{i'} + v_{j'}) \geqslant u_i - u_{i'} > \frac{2}{\ell}$$

where for the inequalities we use the fact that $u_0, \ldots, u_\ell$ and $v_0, \ldots, v_\ell$ are increasing sequences in which the difference between any pair of elements is strictly more than $2/\ell$. This contradicts (4) above. In particular, this implies that $j < j'$. By (4), this yields

$$|(u_i - u_{i'}) - (v_{j'} - v_j)| = \left| \sum_{p=i'+1}^{i} x_k - \sum_{q=j+1}^{j'} y_j \right| \leqslant \frac{2}{\ell}.$$

This implies that to get a set T satisfying the requirements of the lemma, it is sufficient to take $T = \{x_{i'+1}, \ldots, x_i, -y_{j+1}, \ldots, -y_{j'}\}$. Note that $|T| \leqslant 2\ell \leqslant t$, and by the above computation $|\mathrm{Sum}(T)| \leqslant 2/\ell \leqslant 4/(t-1)$. ◀

3 The Lower Bound technique and Tree bias

In this section we will show that tight bounds on the tree bias yield the best possible bound on the relative-rank of set-multilinear low-depth formulas. Specifically, we prove Theorem 3.

3.1 Set-multilinear formulas and Unique Parse Trees

First, it will be helpful to make some structural changes to the formula. We will write a set-multilinear formula as a *small* sum of set-multilinear formulas such that each formula has a unique *parse tree*. In order to describe this we introduce some definitions.

▶ **Definition 10** (Parse Formula). *Let F be a set-multilinear formula. A parse formula F' is obtained from F as follows.*

- *The root $+$ gate is added to F'.*
- *For every $+$ gate added to F', one of its children is added to F'.*
- *For every $\times$ gate added to F', all its children are added to F'.*

Note that, such a parse formula computes a set-multilinear monomial. The polynomial computed by F is the sum of monomials computed by its parse formulas.

Parse trees and W-trees

Let F' be a parse formula from a set-multilinear formula F. We define the parse trees of F as follows. Let g be a $+$ gate with the parent u and child v. We draw a direct edge between u and v and remove the $+$ gate from F'. We do this *short-circuiting* step for each $+$ gate of the parse formula. Similarly, we remove the $+$ root of F'. Let $\mathfrak{T}$ be the tree thus obtained. We call this the *shape* of F'.

Let ℓ be a leaf of $\mathfrak{T}$. It corresponds to a gate g in F which is either a + gate in F' or a leaf in F'. The polynomial computed by g is a linear polynomial on variable set X_i for some $i \in [d]$. We label ℓ with (i, α_i). This way, we label each leaf of $\mathfrak{T}$ with elements of W. We call the W-tree T thus obtained a *parse tree* of F. Note that the depth of T is the same as the product-depth of F.

▶ **Definition 11** (UPT formula). *We say that a set-multilinear formula F is a Unique Parse Tree formula (or UPT) if all the parse trees of F are identical.*

▶ **Lemma 12.** *Let F be a set-multilinear formula of size s and depth Δ. Then F can be written as a sum of at most d^{3d} many set-multilinear UPT formulas such that each formula has size at most s and depth Δ.*

We defer the proof of this lemma to the full version due to lack of space.

3.2 Tree bias lower bounds imply formula lower bounds

In this section, we show how lower bounds on $\mathrm{Treebias}_\Delta(W)$ imply set-multilinear formula lower bounds in the low-degree setting. By Lemma 7, this implies lower bounds for general formulas as well.

We first show this connection for a UPT formula and then use the lemma from the previous section to conclude the same for general set-multilinear formulas. Specifically, we prove the following statement.

▶ **Lemma 13.** *Let n, d be positive integers. Let $\Delta \geqslant 1$. Let W be a multiset of $[-1, 1]$ of size d. Let F be a set-multilinear UPT formula of size s, product-depth Δ, and parse tree T. Assume, moreover, that* $\mathrm{Pathbias}_W(T) = p$. *Then,*

$$\mathrm{relrk}_W(F) \leqslant (s \cdot n^{-p/2}) \cdot n^{-|\mathrm{Sum}(W)|/2}.$$

We first use this lemma to prove part (1) of Theorem 3.

Proof of Part (1) of Theorem 3. Let W and t be as in the statement of Theorem 3. Let F be a set-multilinear formula of product depth Δ and size at most s. From Lemma 12 we know that F can be written as a sum of UPT formulas, say $\Psi_1, \Psi_2, \ldots, \Psi_r$, where $r \leqslant d^{3d}$. We also know that the size of each Ψ_i is at most s and their depth is Δ. Let $\Gamma_1, \ldots, \Gamma_r$ be the parse trees of these formulas and let $p_i = \mathrm{Pathbias}_W(\Gamma_i)$ for $i \in [r]$.

By Lemma 13, for each $i \in [r]$, $\mathrm{relrk}_W(\Psi_i) \leqslant (s \cdot n^{-p_i/2}) \cdot n^{-|\mathrm{Sum}(W)|/2}$. As $\mathrm{Treebias}_\Delta(W) = t$, we have $p_i \geqslant t$ for each $i \in [r]$. Therefore, we get

$$i \in [r], \mathrm{relrk}_W(\Psi_i) \leqslant (s \cdot n^{-t/2}) \cdot n^{-|\mathrm{Sum}(W)|/2}.$$

As $F = \sum_{i=1}^{r} \Psi_i$, $r \leqslant d^{3d}$ and by sub-additivity of relrk, we get the claimed bound on the relrk of F, i.e.

$$\mathrm{relrk}_W(F) \leqslant (d^{3d} \cdot s \cdot n^{-t/2}) \cdot n^{-|\mathrm{Sum}(W)|/2}.$$ ◀

We now prove Lemma 13.

Proof of Lemma 13. We prove the statement by induction on the depth of T (which is also the product depth of F).

Base case. Let $F = \sum_i \prod_j F_{i,j}$ be a set-multilinear UPT formula of product-depth $\Delta = 1$. Let T be the W-tree corresponding to F. Let u_0 be the root of T and let $u_1, \ldots, u_d$ be the children of u_0 with labels $(1, \alpha_1), \ldots, (d, \alpha_d)$, respectively.

By sub-additivity and sub-multiplicativity (Lemma 6, Items 2 and 3) of relrk, we can say that $\mathrm{relrk}_W(F) \leqslant \sum_i \prod_j \mathrm{relrk}_{\{\alpha_j\}}(F_{i,j})$. By using the Imbalance bound (Lemma 6 Item 1) on the relative rank of each $F_{i,j}$ we get that

$$\mathrm{relrk}_W(F) \leqslant \sum_i n^{-\sum_j |\alpha_j|/2} = s \cdot n^{-\sum_j |\alpha_j|/2} = sn^{-p/2} n^{-|\mathrm{Sum}(W)|/2}$$

where the last equality follows from Proposition 8. We get the desired bound.

Induction step. Let $F = \sum_i \prod_j F_{i,j}$ be a set-multilinear UPT formula of depth $\Delta > 1$. Let T be the W-tree corresponding to F. Let u_0 be the root of T and let $u_1, \ldots, u_k$ be the children of u_0. Let $T_1, \ldots, T_k$ be the trees rooted at $u_1, \ldots, u_k$ respectively.

As F is a UPT formula, we have that for each $i \neq i'$ and for any $j \in [k]$, the parse tree of $F_{i,j}$ is the same as the parse tree of $F_{i',j}$. Without loss of generality let us say the parse tree of $F_{i,j}$ is T_j for every i.

Also, for T, let us assume without loss of generality that the path bias of T is realised by a path π, where $\pi = u_0 \cdot u_1 \cdot \pi'$, i.e. specifically it passes through u_1. Let p_1 denote $\mathrm{Pathbias}_{W_{u_1}}(T_1)$.

Finally, let $s_{i,j}$ denote the size of the subformula $F_{i,j}$. Note that $\sum_{i,j} s_{i,j} \leqslant s$.

$$\begin{aligned}
\mathrm{relrk}_W(F) &\leqslant \sum_i \mathrm{relrk}_{W_{u_1}}(F_{i,1}) \cdot \prod_{j \geqslant 2} \mathrm{relrk}_{W_{u_j}}(F_{i,j}) && \text{Properties of relrk} \\
&\leqslant \sum_i \mathrm{relrk}_{W_{u_1}}(F_{i,1}) \cdot \prod_{j \geqslant 2} n^{-|\mathrm{Sum}(u_j)|/2} && \text{Trivial bound on relrk} \\
&\leqslant \sum_i \Big((s_{i,1} \cdot n^{-p_1/2}) \cdot n^{-|\mathrm{Sum}(u_1)|/2} \Big) \cdot \prod_{j \geqslant 2} n^{-|\mathrm{Sum}(u_j)|/2} && \text{Induction Hypothesis} \\
&\leqslant \sum_i s_{i,1} \cdot n^{-\left(p_1 + \sum_{j=1}^k \mathrm{Sum}(u_j)\right)/2} \\
&= \sum_i s_{i,1} \cdot n^{-(p + |\mathrm{Sum}(W)|)/2} && \text{Proposition 8} \\
&\leqslant s \cdot n^{-p/2} \cdot n^{-|\mathrm{Sum}(W)|/2}.
\end{aligned}$$ ◀

3.3 Tree bias upper bounds imply formula upper bounds

We now sketch the proof of the second part of Theorem 3. The main idea is an abstraction of a proof from our earlier result [18][13] where we constructed polynomials to show that our lower bound technique was "tight" for certain concrete spaces of lopsided set-multilinear polynomials. In the second part of Theorem 3, we essentially show that the lower bound proved via tree-bias is tight for *all* lopsided spaces.

The main technical result (which generalizes [18, Lemma 26]) is the following, which handles the case where each $|X_i|$ is a power of 2.

[13] More specifically, this result appeared in a later version of the paper that can be found on ECCC.

▶ **Lemma 14.** *Let n, d be growing parameters and Δ any positive integer. Let $W = \{\alpha_1, \ldots, \alpha_d\} \subseteq [-1, 1]$ be a multiset. Assume that $\mathbb{F}_{\mathrm{sm}}[X_1, \ldots, X_d]$ be a lopsided space of set-multilinear polynomials with $|X_i| = n^{|\alpha_i|} = 2^{k_i}$ for non-negative integers $k_1, \ldots, k_d$.*

Let T be any W-tree of depth Δ with $\mathrm{Pathbias}_W(T) = p$. Then, there is a UPT formula F of parse tree T (and hence product-depth Δ) with at most $d \cdot n^{p/2}$ leaves such that $\mathrm{rank}(M_W(F))$ is as large as possible (i.e. equal to either the number of its rows or columns).

We defer the proof of this lemma, as well as its generalization which yields the second part of Theorem 3, to the full version of the paper.

4 Optimal bounds for depth 3 via our technique

This section is devoted to the proof of Theorem 4, which characterizes (up to constant factors) the maximum possible tree bias of a tree of depth 3.

In proving this theorem, it will be useful to consider a variant on the notion of tree bias defined above, that we will call *node bias*. The node bias of W (at any given depth Δ) is equal to the tree bias of W up to a factor of $O(\Delta)$. By Theorem 3, for constant-depths Δ, the node bias also captures the best lower bound that we can hope to prove via our technique.

▶ **Definition 15** (Node bias). *Fix a W-tree T. For an internal node v of T, we define the bias of v, denoted $\mathrm{bias}(v)$, to be $\sum_u |\mathrm{Sum}(u)|$ where the sum runs over the children u of v. The node bias of T, denoted $\mathrm{Nodebias}_W(T)$, is the largest bias of any internal node v of T. Further, the depth-Δ node bias of W, is the minimum node bias of any depth-Δ, W-tree T. This quantity is denoted $\mathrm{Nodebias}_\Delta(W)$.*

The following basic proposition (proof omitted) relates the node bias of W and the tree bias of W.

▶ **Proposition 16.** *For any depth-Δ W-tree T, we have*

$$\mathrm{Nodebias}_W(T) \leqslant \mathrm{Pathbias}_W(T) + |\mathrm{Sum}(W)| \leqslant \Delta \cdot \mathrm{Nodebias}_W(T).$$

In particular, for any multiset $W \subseteq [-1, 1]$ and any depth Δ, we have $\mathrm{Nodebias}_\Delta(W) \leqslant \mathrm{Treebias}_\Delta(W) + |\mathrm{Sum}(W)| \leqslant \Delta \cdot \mathrm{Nodebias}_\Delta(W)$.

4.1 Some simple claims

This section presents several statements about W-trees. As the proofs are simple, we defer them to the full version of the paper.

Given a partition[14] P of the elements of W, we define the *grouping* W' of W to be the multiset obtained by taking the sums of elements of P. Formally,

$$W' = \{\mathrm{Sum}(A) \mid A \in P\}.$$

The following basic lemma shows how to construct a W-tree from trees of its groupings and subsets.

[14] We follow the usual convention that $\varnothing \notin P$.

▶ **Lemma 17.** *Assume that $P = \{W_1, \ldots, W_t\}$ is a partition of W and let W' be the corresponding grouping. Say we have a W'-tree T' of node bias b' and depth Δ' and for each $i \in [t]$, a W_i-tree T_i of depth Δ_i and node bias at most b_i. Then, there is a W-tree T of node bias at most $\max\{b', b_i\}$ and depth at most $\Delta' + \max_{i\in[t]} \Delta_i$.*

Moreover, if each W_i is sign-monochromatic *(i.e., all elements of W_i have the same sign)*[15]*, then there is a W-tree T of depth Δ' and bias b'.*

▶ **Lemma 18** (Preprocessing Lemma)**.** *Let $W \subseteq [-1, 1]$ be any multiset. Then, there is a partition $P = \{W_1, \ldots, W_t\}$ of W such that each W_i is sign-monochromatic (as in Lemma 17), $\mathrm{Sum}(W_i) \in [-1, 1]$ for all $i \in [t]$ and $\mathrm{Sum}(W_i) \in [-1, -1/2] \cup [1/2, 1]$ for each $i \in \{3, \ldots, t\}$.*

In particular, there is a grouping $W' \subseteq [-1, 1]$ of W such that $|W' \setminus ([-1, -1/2] \cup [1/2, 1])| \leqslant 2$ and for each W'-tree T', there is a W-tree T of same depth and node bias.

The next lemma shows, in particular, how to construct W-trees of depth Δ and node bias $O(d^{1/\Delta})$ for any multiset $W \subseteq [-1, 1]$ of size d and of sum at most 1.

▶ **Lemma 19.** *Let $W \subseteq [-1, 1]$ such that $\|W\|_1 \leqslant L$ and $|\mathrm{Sum}(W)| \leqslant 1$. Then, for any $\Delta \geqslant 1$, there is a W-tree of depth at most Δ and node bias at most $5L^{1/\Delta}$.*

We also have the following simple "pasting" lemma.

▶ **Lemma 20.** *Let $P = \{W_1, \ldots, W_r\}$ be a partition of W and assume that for all i there is a W_i-tree T_i of depth at most Δ, node bias at most b_i and such that the root node of each T_i has bias at most b'_i. Then, there is a W-tree T of depth at most Δ and of node bias at most $\max\{b_1, \ldots, b_r, \sum_i b'_i\}$.*

Finally, the following claim will allow us to balance a given subset of W so that removing this subset results in two sets of absolute sum at most 1.

▶ **Lemma 21** (Balancing lemma)**.** *Say $W \subseteq [-1, 1]$ is such that $|\mathrm{Sum}(W)| \leqslant 1$. Let $W' \subseteq W$ be arbitrary. Then there exists $W'' \subseteq W$ such that $W'' \supseteq W'$ and $\|W'' \setminus W'\|_1 \leqslant |\mathrm{Sum}(W')| + 1$ and $|\mathrm{Sum}(W'')|, |\mathrm{Sum}(W \setminus W'')| \leqslant 1$.*

4.2 Depth-3 trees of small bias

The main theorem of this section is the following.

▶ **Theorem 22.** *Let $W \subseteq [-1, 1]$ be any multiset such that $|W| \leqslant d$ and $|\mathrm{Sum}(W)| \leqslant 1$. Then, there is a W-tree T of depth 3 and node bias $O(d^{1/4})$.*

The rest of the section is devoted to the proof of the above theorem. To construct the required W-tree T, we use the following procedure.

1. Preprocessing: By the Preprocessing lemma (Lemma 18), it suffices to consider multisets W such that $|W \setminus ([-1, -1/2] \cup [1/2, 1])| \leqslant 2$.
2. We apply the following procedure to our multiset W.

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[15] Here, we think of 0 as having the same sign as any other number.

Algorithm 1 $\mathcal{A}(W)$.

Assignment $d := |W|$.
Initialization If $d \leqslant 25$ then return the trivial depth-1 W-tree of node bias at most 25.
Phase 1: As long as it is possible, pick pairwise disjoint sets A such that $|A| \leqslant d^{1/4}$ and

$$\frac{|\text{Sum}(A)|}{|A|} \leqslant \frac{12}{d^{1/2}}.$$

When this is no longer possible, let $A_1, \ldots, A_{e_1}$ be the sequence of sets picked and let $W_1' = A_1 \cup A_2 \cdots \cup A_{e_1}$. Using the Balancing lemma, let $W_1 = W_1' \cup \{a_1, \ldots, a_{f_1}\} \subseteq W$ be such that $\sum_{i \leqslant f_1} |a_i| \leqslant |\text{Sum}(W_1')| + 1$ and $|\text{Sum}(W_1)|, |\text{Sum}(W \backslash W_1)| \leqslant 1$.
Construct a W_1-tree T_1 in the following way. Fix the grouping $\tilde{W}_1$ corresponding to the partition $P_1 = \{A_1, \ldots, A_{e_1}, \{a_1\}, \ldots, \{a_{f_1}\}\}$ of W_1. For each element of P_1, construct a trivial tree of depth-1 and for the grouping $\tilde{W}_1$, construct a depth-2 tree $\tilde{T}_1$ of node bias at most $5\sqrt{\|\tilde{W}_1\|_1}$ (using Lemma 19). Combine these using Lemma 17 to get a tree T_1 of depth 3 for W_1.
Set $W' = W \backslash W_1$ and continue.
Phase 2: As long as it is possible, pick pairwise disjoint sets $B \subseteq W'$ such that $|B| \leqslant d^{1/2}$ and

$$\frac{|\text{Sum}(B)|}{|B|} \leqslant \frac{12}{d^{3/4}}.$$

When this is no longer possible, let $B_1, \ldots, B_{e_2}$ be the sequence of sets picked and let $W_2' = B_1 \cup B_2 \cdots \cup B_{e_2}$. Using the balancing lemma, let $W_2 = W_2' \cup \{b_1, \ldots, b_{f_2}\} \subseteq W'$ be such that $\sum_{i \leqslant f_1} |b_i| \leqslant |\text{Sum}(W_2')| + 1$ and $|\text{Sum}(W_2)|, |\text{Sum}(W' \backslash W_2)| \leqslant 1$.
Construct a W_2-tree T_2 in the following way. Fix the grouping $\tilde{W}_2$ corresponding to the partition $P_2 = \{B_1, \ldots, B_{e_2}, \{b_1\}, \ldots, \{b_{f_2}\}\}$ of W_2. Construct a trivial depth-1 $\tilde{W}_2$-tree $\tilde{T}_2$ of node bias $\|\tilde{W}_2\|_1$. For each element B of P_2, construct a depth-2 tree of node bias at most $5\sqrt{\|B\|_1}$ (using Lemma 19). Combine these using Lemma 17 to get a tree T_2 of depth 3 for W_2.
Set $W'' = W' \backslash W_2$ and continue.
Recursive call Compute $T_3 = \mathcal{A}(W'')$.
Return The W-tree T of node bias at most $b_1 + b_2 + b_3$, where $b_i = \text{Nodebias}(T_i)$ (for $i \in [3]$) given by T_1, T_2, T_3 and Lemma 20.

We now analyze the above construction. We first state a technical lemma.

▶ **Lemma 23.** *If $d > 25$, then after Phases 1 and 2, we have $|W''| \leqslant d/2$.*

Let us assume the above lemma for now and prove the theorem.

Let $b_i(d)$ denote the node bias of the tree T_i ($i \in [3]$) assuming that the word W has size at most d. Then, the node bias of the tree is $b_1(d) + b_2(d) + b_3(d)$. By Lemma 23, we can bound $b_3(d)$ by $b_1(d/2) + b_2(d/2) + b_3(d/2)$. Continuing recursively in this way (until d becomes smaller than 25) we have

$$\text{Nodebias}(T) \leqslant \left(\sum_{i \geqslant 0} b_1(d/2^i) + b_2(d/2^i) \right) + 25.$$

So to prove Theorem 22, it suffices to show that $b_1(d), b_2(d) \leqslant O(d^{1/4})$. From now on, we fix d and let $b_i = b_i(d)$ for $i \in [2]$.

We first bound b_1. By construction each element of the partition P_1 is a set A of size at most $d^{1/4}$ and hence has a depth-1 tree of node bias at most $d^{1/4}$. Moreover, we have

$$\begin{aligned}\|\tilde{W}_1\|_1 &= \sum_{i\leqslant e_1} |\mathrm{Sum}(A_i)| + \sum_{j\leqslant f_1} |a_j| \\ &\leqslant \sum_{i\leqslant e_1} |\mathrm{Sum}(A_i)| + |\mathrm{Sum}(W_1')| + 1 \leqslant 2\sum_{i\leqslant e_1} |\mathrm{Sum}(A_i)| + 1 \\ &\leqslant 2\sum_{i\leqslant e_1} \frac{12|A_i|}{d^{1/2}} + 1 \leqslant O(d^{1/2}).\end{aligned}$$

Hence, the tree $\tilde{T}_1$ has node bias $\tilde{b}_1 = O(d^{1/4})$. Hence, by Lemma 17, we see that $b_1 \leqslant O(d^{1/4})$.

We can bound b_2 similarly. By construction, each element of P_2 is a set B of size at most $d^{1/2}$ and hence by Lemma 19 has a depth-2 tree of node bias at most $5d^{1/4}$. Moreover, we have

$$\begin{aligned}\|\tilde{W}_2\|_1 &= \sum_{i\leqslant e_2} |\mathrm{Sum}(B_i)| + \sum_{j\leqslant f_1} |b_j| \\ &\leqslant \sum_{i\leqslant e_2} |\mathrm{Sum}(B_i)| + |\mathrm{Sum}(W_2')| + 1 \leqslant 2\sum_{i\leqslant e_2} |\mathrm{Sum}(B_i)| + 1 \\ &\leqslant 2\sum_{i\leqslant e_1} \frac{12|B_i|}{d^{3/4}} + 1 \leqslant O(d^{1/4}).\end{aligned}$$

In particular, this implies that the tree $\tilde{T}_2$ has node bias $\tilde{b}_2 = O(d^{1/4})$. In particular, by Lemma 17, we see that $b_2 \leqslant O(d^{1/4})$.

Thus, we have shown that $b_1, b_2 = O(d^{1/4})$ and we are done.

It remains only to prove Lemma 23, which we do now.

Proof of Lemma 23. Let $d'' = |W''|$. Assuming that $d > 25$ and $d'' > d/2$, we will show that Phases 1 and 2 of the algorithm could not have concluded, and hence derive a contradiction.

Let W''_+ and W''_- denote the positive and negative elements of W'' respectively. Recall that $W'' \subseteq W$ and the latter set contains at most two elements of absolute value less than 1/2 (by the preprocessing in Step 1). Further using the fact that $|\mathrm{Sum}(W'')| \leqslant 1$, it is easy to see that $|W''_+|, |W''_-| \geqslant d'''$ where $d''' = (d'' - 4)/3 > 2$.

By Lemma 9, it follows that there is a non-empty set $T \subseteq W''$ of size $t \leqslant \sqrt{d'''} + 1$ such that $|\mathrm{Sum}T| \leqslant 4/\sqrt{d'''}$. Since $d \geqslant 24$, this set T has size at most $\sqrt{d}$ and satisfies $|\mathrm{Sum}T| \leqslant 12/\sqrt{d}$.

Now, we do a short case analysis. Assume $|T| \leqslant d^{1/4}$. Then, T is the kind of set that the algorithm tries to find in Phase 1. Hence, the existence of such a T tells us that Phase 1 could not have concluded.

Otherwise, we have $|T| > d^{1/4}$. In this case, we have $|\mathrm{Sum}T|/|T| \leqslant 12d^{-3/4}$ and is hence the kind of set that the algorithm tries to find in Phase 2. Hence, the existence of such a T tells us that Phase 2 could not have concluded.

In either case, we are done. ◀

4.3 Optimality of the quartic bound

We will show here that the bound of Theorem 22 is optimal.

▶ **Proposition 24.** *Let d be a growing integer parameter. There exists a multiset $W \subseteq [-1, 1]$ such that $|W| \leqslant d$, $|\mathrm{Sum}W| \leqslant 1$, and for all W-tree T of depth 3, T has node bias at least $\Omega(d^{1/4})$.*

Proof. If $d < 16$ the result follows immediately (just adapt the constant in the $\Omega()$ to deal with these cases). So let us assume that $d \geqslant 16$. Let d' be the largest integer such that $d' \leqslant d$ and d' is a fourth power of an integer. So $d'^{1/4} \geqslant 2$ and $d' \geqslant d/16$.

Let q be the closest integer to $\frac{d'}{2-1/d'^{1/4}+1/(2d'^{3/4})}$. So $\left|q - \frac{d'}{2-1/d'^{1/4}+1/(2d'^{3/4})}\right| \leqslant \frac{1}{2}$. Let us construct W with q copies of $1 - 1/d'^{1/4} + 1/(2d'^{3/4})$ and $p = d' - q$ copies of -1. So $|W| \leqslant d' \leqslant d$ and

$$\begin{aligned}|\mathrm{Sum}(W)| &= \left|-p + q(1 - 1/d'^{1/4} + 1/(2d'^{3/4}))\right| \\ &= \left|-d' + q(2 - 1/d'^{1/4} + 1/(2d'^{3/4}))\right| \\ &\leqslant \left|(-d' + d')\right| + \frac{1}{2}\left|(2 - 1/d'^{1/4} + 1/(2d'^{3/4}))\right| \leqslant 1.\end{aligned}$$

It is sufficient to prove that any W-tree has large enough node bias.

Let T be any W-tree. Let us assume that $\mathrm{Nodebias}_W(T) < d'^{1/4}/4$.

Since every internal node α at distance two of the roots with k children (in particular the children of α are leaves of T) has bias at least $\mathrm{bias}(\alpha) \geqslant k \min_{v\in W} |v| \geqslant k/2$, it implies that $k < d'^{1/4}/2$.

Assume then that there is an internal node α at distance one of the root such that the subtree rooted in α has at least $d'^{1/2}$ leaves. Notice that for any children β of α with p_β negative children and q_β positive ones we have

$$\begin{aligned}|\mathrm{Sum}\beta| &= \left|-p_\beta + q_\beta(1 - 1/d'^{1/4} + 1/(2d'^{3/4}))\right| \\ &= \left|(-p_\beta + q_\beta) - q_\beta(1 - 1/2d'^{1/2})/d'^{1/4}\right|.\end{aligned}$$

Since $q_\beta \leqslant p_\beta + q_\beta < d'^{1/4}/2$ and p_β, q_β are integers, it implies that the fractional part of $|\mathrm{Sum}\beta|$ is at least $q_\beta(1 - 1/2d'^{1/2})/d'^{1/4}$. Moreover, if $|\mathrm{Sum}(\beta)| < 1$, it means that $q_\beta \geqslant p_\beta$, i.e., $q_\beta \geqslant (p_\beta + q_\beta)/2$. Hence in all cases,

$$|\mathrm{Sum}(\beta)| \geqslant \frac{q_\beta + p_\beta}{2} \cdot \frac{1}{2} \cdot \frac{1}{d'^{1/4}}.$$

Consequently, $\mathrm{bias}(\alpha) = \sum_{\beta \text{ child of } \alpha} |\mathrm{Sum}(\beta)| \geqslant \frac{d'^{1/2}}{4d'^{1/4}} = \frac{d'^{1/4}}{4}$, which contradicts the hypothesis. So any node at depth 1 of the tree has less than $d'^{1/2}$ leaves in its subtree.

Let us show that finally the root ρ of T has large bias. Let β one of its children. Say that in the tree rooted in β, there are p_β negative leaves and q_β positive ones. So,

$$\begin{aligned}|\mathrm{Sum}\beta| &= \left|-p_\beta + q_\beta(1 - 1/d'^{1/4} + 1/(2d'^{3/4}))\right| \\ &= \left|(-p_\beta d'^{1/4} + q_\beta d'^{1/4} - q_\beta)\frac{1}{d'^{1/4}} + q_\beta \frac{1}{2d'^{3/4}}\right|.\end{aligned}$$

Since $q_\beta/(2d'^{3/4}) < 1/(2d'^{1/4})$, it implies that the distance of $|\mathrm{Sum}(\beta)|$ to the set $\mathbb{N}/d'^{1/4}$ is at lest $q_\beta/(2d'^{3/4})$. Again, if $|\mathrm{Sum}(\beta)| < 1$, it ensures that $q_\beta \geqslant (p_\beta + q_\beta)/2$. So in all cases,

$$|\mathrm{Sum}(\beta)| \geqslant \frac{p_\beta + q_\beta}{2} . \frac{1}{2d'^{3/4}}.$$

Consequently,

$$\mathrm{bias}(\rho) = \sum_{\beta \text{ child of } \rho} |\mathrm{Sum}(\beta)| \geqslant \frac{1}{4d'^{3/4}} \sum_{\beta \text{ child of } \rho} p_\beta + q_\beta = \frac{d'^{1/4}}{4}$$

which again contradicts the hypothesis.

In conclusion, we have that for any W-tree T and $\mathrm{Nodebias}_W(T) \geqslant \frac{d'^{1/4}}{4} \geqslant \frac{d^{1/4}}{8}$. ◀

▶ Remark 25. We can generalize the previous proof to larger depths by defining q to be the closest integer to $d'/(2 + \sum_{i=1}^{\Delta-1}(-1)^i/d'^{(2^i-1)/2^{\Delta-1}})$. It implies that for all Δ, there exists a multiset W such that any W-tree of depth Δ has node bias at least $\Omega(d^{1/2^{\Delta-1}})$. It improves the constant in the exponent slightly in the lower bound from [18].

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