On the Efficient Implementation of High Accuracy Optimality of Profile Maximum Likelihood

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Abstract

We provide an efficient unified plug-in approach for estimating symmetric properties of distributions given $n$ independent samples. Our estimator is based on profile-maximum-likelihood (PML) and is sample optimal for estimating various symmetric properties when the estimation error $\epsilon \gg n^{-1/3}$. This result improves upon the previous best accuracy threshold of $\epsilon \gg n^{-1/4}$ achievable by polynomial time computable PML-based universal estimators [ACSS21][ACSS20]. Our estimator reaches a theoretical limit for universal symmetric property estimation as [Han21] shows that a broad class of universal estimators (containing many well known approaches including ours) cannot be sample optimal for every 1-Lipschitz property when $\epsilon \ll n^{-1/3}$.

1 Introduction

Given $n$ independent samples $y_1, ..., y_n \in D$ from an unknown discrete distribution $p \in \Delta^D$ the problem of estimating properties of $p$, e.g., entropy, distance to uniformity, support size and coverage are among the most fundamental in statistics and learning. Further, the problem of estimating symmetric properties of distributions $p$ (i.e., properties invariant to label permutations) are well studied and have numerous applications [Cha84][BF93][CCG12][TE87][Für05][KLR99][PBG01][DS13][RCS09][GTPB07][HHRB01].

Over the past decade, symmetric property estimation has been studied extensively and there have been many improvements to the time and sample complexity for estimating different properties, e.g. support [VV11b][WY15], coverage [ZVV16][OSW16], entropy [VV11b][WY16][VHW15], and distance to uniformity [VV11a][JHW16]. Towards unifying the attainment of computational-efficiency, sample-optimal estimators a striking work of [ADOS17] provided a universal plug-in approach based on a (approximate) profile maximum likelihood (PML) distribution, that (approximately) maximizes the likelihood of the observed profile (i.e. multiset of observed frequencies).

Formally, [ADOS17] showed that given $y_1, ..., y_n$, if there exists an estimator for a symmetric property $f$ achieving accuracy $\epsilon$ and failure probability $\delta$, then this PML-based plug-in approach achieves error $2\epsilon$ with failure probability $\delta \exp (3\sqrt{n})$. As the failure probability $\delta$ for many estimators for well-known properties (e.g., support size and coverage, entropy, and distance to uniformity) is roughly $\exp (-\epsilon^2 n)$, this result implied a sample optimal unified approach for estimating these properties when the estimation error $\epsilon \gg n^{-1/4}$.

This result of [ADOS17] laid the groundwork for a line of work on the study of computational and statistical aspects of PML-based approaches to symmetric property estimation. For example, follow up work of [HS21] improved the analysis of [ADOS17] and showed that the failure probability of PML is at most \( \delta^{1-c} \exp(-n^{1/3}+c) \), for any constant \( c > 0 \) and therefore it is sample optimal in the regime \( \epsilon \gg n^{-1/3} \). The condition \( \epsilon \gg n^{-1/3} \) on the optimality of PML is tight [Han21], in the sense that, PML is known to be not sample optimal in the regime \( \epsilon \ll n^{-1/3} \). In fact, no estimator (that obeys some mild conditions), is sample optimal for estimating all symmetric properties in the regime \( \epsilon \ll n^{-1/3} \); see Section 2 after Theorem 2.6 for details.

We also remark that the statistical guarantees in [ADOS17, HS21] hold for any \( \beta \)-approximate PML\(^1\) for suitable values of \( \beta \). In particular, [HS21] showed that any \( \beta \)-approximate PML for \( \beta > \exp(-n^{1-c'}) \) and any constant \( c' > 0 \), has a failure probability of \( \delta^{1-c} \exp(n^{1/3+c'}+n^{1-c'}) \) for any constant \( c > 0 \). These results further imply a sample optimal estimator in the regime \( \epsilon \gg n^{-\min(1/3,c'/2)} \) for properties with failure probability less than \( \exp(-\epsilon^2n) \). Note that better approximation leads to a larger range of \( \epsilon \) for which the estimator is sample optimal.

Regarding computational aspects of PML, [CSS19a] provided the first efficient algorithm with a non-trivial approximation guarantee of \( \exp(-n^{2/3}\log n) \), which further implied a sample optimal universal estimator for \( \epsilon \gg n^{-1/6} \). This was then improved by [ACSS21] which showed how to efficiently compute PML to higher accuracy of \( \exp(-\sqrt{n}\log n) \) thereby achieving a sample optimal universal estimator in the regime \( \epsilon \gg n^{-1/4} \). The current best polynomial time approximate PML algorithm by [ACSS20] achieves an accuracy of \( \exp(-k\log n) \), where \( k \) is the number of distinct observed frequencies. Although this result achieves better instance based statistical guarantees, in the worst case it still only implies a sample optimal universal estimator in the regime \( \epsilon \gg n^{-1/4} \).

In light of these results, a key open problem is to close the gap between the regimes \( \epsilon \gg n^{-1/3} \) and \( \epsilon \gg n^{-1/4} \), where the former is the regime in which PML based estimators are statistically optimal and the later is the regime where efficient PML based estimators exist. In this work we ask:

**Is there an efficient approximate PML-based estimator that is sample optimal for \( \epsilon \gg n^{-1/3} \)?**

In this paper, we answer this question in the affirmative. In particular, we give an efficient PML-based estimator that has failure probability at most \( \delta^{1-c} \exp(n^{1/3+c}+n^{1-c'}) \), and consequently is sample optimal in the regime \( \epsilon \gg n^{-1/3} \). As remarked, this result is tight in the sense that PML and a broad class of estimators are known to be not optimal in the regime \( \epsilon \ll n^{-1/3} \).

To obtain this result we depart slightly from the previous approaches in [ADOS17, CSS19a, ACSS21]. Rather than directly compute an approximate PML distribution we compute a weaker notion of approximation which we show suffices to get us the desired universal estimator. We propose a notion of a \( \beta \)-weak approximate PML distribution inspired by [HS21] and show that an \( \exp(-n^{1/3}\log n) \)-weak approximate PML achieves the desired failure probability of \( \delta^{1-c} \exp(n^{1/3+c'}) \) for any constant \( c > 0 \). Further, we provide an efficient algorithm to compute an \( \exp(-n^{1/3}\log n) \)-weak approximate PML distribution. Our paper can be viewed as an efficient algorithmic instantiation of [HS21].

Ultimately, our algorithms use the convex relaxation presented in [CSS19a] and provide a new rounding algorithm. We differ from the previous best \( \exp(-k\log n) \) approximate PML algorithm [ACSS20] only in the matrix rounding procedure which controls the approximation guarantee. At a high level, the approximation guarantee for the rounding procedure in [ACSS20] is exponential in the sum of matrix dimensions. In the present work, we need to round a rectangular matrix with an approximation exponential in the smaller dimension, which may be infeasible for arbitrary matrices. Our key technical innovation is to introduce a swap operation (see Section 4.1) which facilitates such an approximation guarantee. In addition to a better approximation guarantee than [ACSS20], our algorithm also exhibits better run times (see Section 3).

**Organization:** We introduce preliminaries in Section 1.1. In Section 2 we state our main results and also cover related work. In Section 3, we provide the convex relaxation to PML studied in [CSS19a, ACSS21]. Finally, in Section 4, we provide a proof sketch of our main computational result. Many proofs are then deferred to the appendix.

\(^1\beta\)-approximate PML is a distribution that achieves a multiplicative \( \beta \)-approximation to the PML objective.
1.1 Preliminaries

General notation: For matrices $A, B \in \mathbb{R}^{s \times t}$, we use $A \preceq B$ to denote that $A_{ij} \leq B_{ij}$ for all $i \in [s]$ and $j \in [t]$. We let $[a, b]$ and $[a, b]_{\mathbb{Z}}$ denote the interval $[a, b]$ of reals and reals respectively. We use $O(\cdot)$, $\Omega(\cdot)$ notation to hide all polylogarithmic factors in $n$ and $N$. We let $a_n \gg b_n$ to denote that $a_n \in O(b_n n^c)$ or $b_n \in O(n^{-c} a_n)$, for some small constant $c > 0$.

Throughout this paper, we assume we receive a sequence of $n$ independent samples from a distribution $p \in \Delta^D$, where $\Delta^D \overset{\text{def}}{=} \{q \in [0, 1]^D \ | \ \|q\|_1 = 1\}$ is the set of all discrete distributions supported on domain $D$. Let $\Delta^n$ be the set of all length $n$ sequences of elements of $D$ and for $y^n \in \Delta^n$ let $y_i^n$ denoting its $i$th element. Let $f(y^n, x) \overset{\text{def}}{=} |\{i \in [n] \ | \ y_i^n = x\}|$ and $p_x$ be the frequency and probability of $x \in D$ respectively. For a sequence $y^n \in \Delta^n$, let $M = \{f(y^n, x)\}_{x \in D} \setminus \{0\}$ be the set of all its non-zero distinct frequencies and $m_1, m_2, \ldots, m_M$ be these distinct frequencies. The profile of a sequence $y^n$ denoted $\phi = \Phi(y^n)$ is a vector in $\mathbb{Z}^M$, where $\phi_j \overset{\text{def}}{=} |\{x \in D \ | \ f(y^n, x) = m_j\}|$ is the number of domain elements with frequency $m_j$. We call $n$ the length of profile $\phi$ and let $\Phi^n$ denote the set of all profiles of length $n$. The probability of observing sequence $y^n$ and profile $\phi$ for distribution $p$ are $\mathbb{P}(p, y^n) = \prod_{x \in D} p_x^{f(y^n, x)}$ and $\mathbb{P}(p, \phi) = \sum_{\{y^n \in \Delta^n \ | \ \Phi(y^n) = \phi\}} \mathbb{P}(p, y^n)$.

Profile maximum likelihood: A distribution $p_\phi \in \Delta^D$ is a profile maximum likelihood (PML) distribution for profile $\phi \in \Phi^n$ if $p_\phi \overset{\text{def}}{=} \text{argmax}_{p \in \Delta^n} \mathbb{P}(p, \phi)$. Further, a distribution $p_\beta$ is a $\beta$-approximate PML distribution if $\mathbb{P}(p_\beta, \phi) \geq \beta \cdot \mathbb{P}(p_\phi, \phi)$. For a distribution $p$ and a length $n$, let $X$ be a random variable that takes value $\phi \in \Phi^n$ with probability $\mathbb{P}(p, \phi)$. We call $H(X)$ (entropy of $X$) the profile entropy with respect to $p$ and denote it by $H(\Phi^n, p)$.

Probability discretization: Let $R \overset{\text{def}}{=} \{r_i\}_{i \in [1, \ell]}$ be a finite discretization of the probability space, where $r_i \in [0, 1]_{\mathbb{R}}$ and $\ell \overset{\text{def}}{=} |R|$. We call $q \in [0, 1]_{\mathbb{R}}^\ell$ a pseudo-distribution if $\|q\|_1 \leq 1$ and a discrete pseudo-distribution with respect to $R$ if all its entries are in $R$ as well. We use $\Delta^\ell_{\text{pseudo}}$ and $\Delta^\ell_{\text{disc}}$ to denote the set of all pseudo-distributions and discrete pseudo-distributions with respect to $R$ respectively. In our work, we use the following most commonly used [CSS19a, ACSS21, ACSS20] probability discretization set. For any $\alpha > 0$,

$$R_{n, \alpha} \overset{\text{def}}{=} \{1\} \cup \left\{ \frac{1}{2n^2(1 + n^{-\alpha})^i} \ | \ \text{for all } i \in \mathbb{Z}_{\geq 0} \text{ such that } \frac{1}{2n^2(1 + n^{-\alpha})^i} \leq 1 \right\}. \quad (1)$$

For all probability terms defined involving distributions $p$, we extend those definitions to pseudo distributions $q$ by replacing $p_x$ with $q_x$ everywhere.

See Appendix A for the definition of an estimator and optimal sample complexity.

2 Results

Here we provide our main results. In our first result (Theorem 2.2), we show that a weaker notion of approximate PML suffices to obtain the desired universal estimator. Later we show that these weaker approximate PML distributions can be efficiently computed (Theorem 2.3).

Definition 2.1. Given a profile $\phi$, we call a distribution $p' \in \Delta^D$ a $\beta$-approximate PML distribution with respect to $R$ if $\mathbb{P}(p', \phi) \geq \beta \cdot \max_{q \in \Delta^\ell_{\text{disc}}} \mathbb{P}\left(\frac{q}{\|q\|_1}, \phi\right)$.

The above definition generalizes $\beta$-approximate PML distributions which is simply the special case when $R = [0, 1]_{\mathbb{R}}$. Using our new definition, we show that for a specific choice of the discretization set $R_{n, 1/3}$, a distribution $p'$ that is an approximate PML with respect to $R_{n, 1/3}$ suffices to obtain a universal estimator; this result is formally stated below.

Theorem 2.2 (Competitiveness of an approximate PML w.r.t $R$). For symmetric property $f$, suppose there exists an estimator $\hat{f}$ that takes input a profile $\phi \in \Phi^n$ drawn from $p \in \Delta^D$ and satisfies,

$$\mathbb{P}\left(\|f(p) - \hat{f}(\phi)\| \geq \epsilon\right) \leq \delta,$$
then for $R = R_{n, 1/3}$ (See Equation (1)), a discrete pseudo distribution $q' \in \Delta^ω_2$ such that $q' / \|q'\|_1$ is an $\exp(-O(|R| \log n))$-approximate PML distribution with respect to $R$ satisfies,

$$\mathbb{P}\left( \left| f\left( \frac{q'}{\|q'\|_1} \right) - f(p) \right| \geq 2e \right) \leq \delta^{1-c} \exp(O(n^{1/3+c})) , \quad \text{for any constant } c > 0 . \quad (2)$$

The proof of the above theorem is implicit in the analysis of [HS21], however we provide a short simpler proof using their continuity lemma (Lemma 2 in [HS21]). Note that the bound on the failure probability we get is the same asymptotically as that of exact PML from [HS21], which is known to be tight [Han21]. Furthermore, to achieve such an improved failure probability bound all we need is an approximate PML distribution with respect to $R$, for some $R$ which is of small size. Taking advantage of this fact and building upon [CSS19a, ACSS21], we provide a new rounding algorithm that outputs the desired approximate PML distribution with respect to $R$.

**Theorem 2.3 (Computation of an approximate PML w.r.t $R$).** We provide an algorithm that given a probability discretization set $R = R_{n, \alpha}$ for $\alpha > 0$ (See Equation (1)) and a profile $\phi$ with $k$ distinct frequencies, runs in time $\tilde{O}\left(|R| + \frac{n}{\min(k, |R|)} \left( \min(|R|, n/k)\omega + \min(|R|, k^2) \right) \right)$, where $\omega < 2.373$ is the current matrix multiplication constant [Wil12, Gal14, AW21] and returns a pseudo distribution $q' \in \Delta^ω_2$ such that,

$$\mathbb{P}\left( \frac{q'}{\|q'\|_1} , \phi \right) \geq \exp(-O(\min(k, |R|) \log n)) \cdot \max_{q \in \Delta^ω_2} \mathbb{P}\left( \frac{q}{\|q\|_1} , \phi \right).$$

When $R = R_{n, 1}$, our algorithm computes an $\exp(-O(k \log n))$ approximate PML distribution, therefore our result is at least as good as the previous best known approximate PML algorithm due to [ACSS20]. In comparison to [ACSS20], our rounding algorithm is simpler and we suspect, more practical. We provide a more detailed comparison to it later in this section.

**Applications:** Our main results have several applications which we discuss here. First note that, combining Theorem 2.2 and 2.3 immediately yields the following corollary.

**Corollary 2.4 (Efficient unified estimator).** Given a profile $\phi \in \Phi^n$ with $k$ distinct frequencies, we can compute an approximate PML distribution $q'$ that satisfies Equation (2) in Theorem 2.2 in time $\tilde{O}\left(\frac{n}{\min(k, n/k)} \left( \min(n^{1/3}, n/k)\omega + \min(n^{1/3}, n/k)^2 \right) \right)$.

For many symmetric properties the failure probability is exponentially small as stated below.

**Lemma 2.5 (Lemma 2 in [ADOS17], Theorem 3 in [HS21]).** For distance to uniformity, entropy, support size and coverage, and sorted $\ell_1$ distance there exists an estimator that is sample optimal and the failure probability is at most $\exp(-c^2 n^{1-\alpha})$ for any constant $\alpha > 0$.

The above result combined with Corollary 2.4 immediately yields the following theorem.

**Theorem 2.6 (Efficient sample optimal unified estimator).** There exists an efficient approximate PML-based estimator that for $\epsilon \gg n^{-1/3}$ and symmetric properties such as, distance to uniformity, entropy, support size and coverage, and sorted $\ell_1$ distance achieves optimal sample complexity and has failure probability upper bounded by $\exp(-n^{1/3})$.

As our work computes an $\exp(-O(k \log n))$ approximate PML, we recover efficient version of Lemma 2.3 and Theorem 2.4 from [ACSS20]. The first result uses $\exp(-O(k \log n))$ approximate PML algorithm to efficiently implement an estimator that has better statistical guarantees based on profile entropy [HO20] (See Section 1.1). The second result provides an efficient implementation of the PseudoPML estimators [CSS19, HO19]. Please refer to the respective papers for further details.

**Tightness of our result:** Recall that [HS21] showed that the failure probability of an (approximate) PML based estimator is upper bounded by $\delta^{1-c} \exp(-n^{1/3+c})$, for any constant $c > 0$. This result further implied a sample optimal universal estimator in the regime $\epsilon \gg n^{-1/3}$ for various symmetric properties (Theorem 2.6). In our work, we efficiently recover these results and a natural question to ask here is if these results can be improved.

As remarked earlier, [Han21] showed that the condition for optimality of PML ($\epsilon \gg n^{-1/3}$) is in some sense tight. More formally, they showed that PML is not sample optimal in estimating every
where the terms correspond to the time required to solve the convex program, sparsify and round
\[ \exp(\text{loss}) \]
with a better approximation loss of \( k \).

A convex approximation to PML was first provided in [CSS19a] and a better analysis for it is shown in
[ACSS21]. In particular, [CSS19a] and [ACSS21] showed that an integral optimal solution to step 1
approximates the PML up to accuracy \( \exp(-n^{2/3} \log n) \) and \( \exp(-\min(k, |R|) \log n) \) respectively,
where \( k \) and \( |R| \) are the number of distinct frequencies and distinct probability values respectively.
In addition to the loss from convex approximation, the previous algorithms also incurred a loss in
the rounding step (Step 2). The loss in the rounding step for the previous works is bounded by
\[ \exp(-n^{2/3} \log n) \] [CSS19a], \( \exp(-\sqrt{n} \log n) \) [ACSS21] and \( \exp(-k \log n) \) [ACSS20].

In our work, we show that there exists a choice of \( R (= R_{n,1/3}) \) that is of small size \( (|R| \leq n^{1/3}) \)
and suffices to get the desired universal estimator. As \( |R| \leq n^{1/3} \), our approach only incurs a loss of
\( \exp(-\min(k, |R|) \log n) \in \exp(-n^{1/3} \log n) \) in the convex approximation step (Step 1). Furthermore
for the rounding step (Step 2), we provide a new simpler and a practical rounding algorithm
with a better approximation loss of \( \exp(-O(\min(k, |R|) \log n)) \in \exp(-O(n^{-1/3} \log n)) \).

Regarding the run times, both [ACSS20] and ours have run times of the form \( T_{\text{solve}} + T_{\text{sparsify}} + T_{\text{round}} \),
where the terms correspond to the time required to solve the convex program, sparsify and round
a solution. In our algorithm, we pay the same cost as [ACSS20] for the first two steps but our run
time guarantees are superior to theirs in the rounding step. In particular, the run time of [ACSS20]
is shown as a large polynomial and perhaps not practical as their approach requires enumerating all the
approximate min cuts. In contrast, our algorithm has a run time that is subquadratic.

Other related work  PML was introduced by [OSS+04]. Many heuristic approaches have been proposed
to compute approximate PML, such as the EM algorithm in [OSS+04], an algebraic approaches
in [ADM10]. Bethe approximation in [Von12] and [Von14], and a dynamic programming approach
in [PJW17]. For the broad applicability of PML in property testing and to estimate other symmetric
properties please refer to [HO19]. Please refer to [HO20] for details related to profile entropy. Other
approaches for designing universal estimators are: [VV11b] based on [ET76], [HJW18] based on
local moment matching, and variants of PML by [CSS19b, HO19] that weakly depend on the target
property that we wish to estimate. Optimal sample complexities for estimating many symmetric prop-
erties were also obtained by constructing property specific estimators, e.g. support [VV11b, WY15],
support coverage [OSW16, ZVV+16], entropy [VV11b, WY16, IJHW15], distance to uniformity [VV11a, IJHW16], sorted \( \ell_1 \) distance [VV11a, IJHW18], Renyi entropy [AOST14, AOST17],
KL divergence [BZLV16, IJHW16] and others.

Limitations of our work One of the limitations of all the provable approximate PML algo-
rithms [CSS19a, ACSS21, ACSS20] (including ours) is that they require the solution of a convex
program that approximates the PML objective and all these previous works use the CVX solver which
is not practical for large sample instances; note that our results hold for small error regimes which
lead to such large sample instances. Therefore, designing a practical algorithm to solve the convex
program is an important future research direction. As discussed above, local moment matching
(LLM) based approach is another universal approach for property estimation. It is unclear which of
the two (PML or LLM) can lead to practical algorithms.
3 Convex relaxation to PML

Here we restate the convex program from [CSS19a] that approximates the PML objective. The current best analysis of this convex program is in [ACSS21]. We first describe the notation and later state several results from [CSS19a, ACSS21] that capture the guarantees of the convex program.

Notation: For any matrices $X \in \mathbb{R}^{a \times c}$ and $Y \in \mathbb{R}^{b \times c}$, we let $\text{concat}(X, Y)$ denote the matrix $W \in \mathbb{R}^{(a+b) \times c}$ that satisfies, $W_{i,j} = X_{i,j}$ for all $i \in [a]$ and $j \in [c]$ and $W_{a+i,j} = Y_{ij}$ for all $i \in [b]$ and $j \in [c]$. Recall we let $R \equiv \{ r_i \}_{i \in [c]}$ be a finite discretization of the probability space, where $r_i \in [0, 1]^c$ and $\ell \equiv |R|$. Let $r_i \in [0, 1]^c$ be a vector whose $i$’th element is equal to $r_i$.

Lemma 3.1 (Lemma 4.4 in [CSS19a]). Let $R = R_{n,\alpha}$ for some $\alpha > 0$. For any profile $\phi \in \Phi^n$ and distribution $p \in \Delta^P$, there exists a pseudo distribution $q \in \Delta^P_R$ that satisfies $\mathbb{P}(p, \phi) \geq \mathbb{P}(q, \phi) \geq \exp(-\alpha n - 6) \max_{p \in \Delta^P} \mathbb{P}(p, \phi)$.

For any probability discretization set $R$, profile $\phi$ and pseudo distribution $q \in \Delta^P_R$, define:

$Z^n_R \equiv \{ X \in \mathbb{R}^{c \times [0,k]} \mid X_{11} \in \mathbb{Z}^c, [X^\top 1]_j = \phi_j \text{ for all } j \in [1, k] \text{ and } r^\top X1 \leq 1 \}$, \hspace{1cm} (3)

$Z^n_R \text{frac} \equiv \{ X \in \mathbb{R}^{c \times [0,k]} \mid [X^\top 1]_j = \phi_j \text{ for all } j \in [1, k] \text{ and } r^\top X1 \leq 1 \}$, \hspace{1cm} (4)

The $j$’th column corresponds to frequency $m_j$ and we use $m_0 \equiv 0$ to capture the unseen elements. Without loss of generality, we assume $m_0 < m_1 < \cdots < m_k$. Let $C_{ij} \equiv m_j \log r_i$ for all $i \in [\ell]$ and $j \in [0,k]$. The objective of the optimization problem is: for any $X \in \mathbb{R}^{c \times [0,k]}$, define,

$g(X) \equiv \exp \left( \sum_{i \in \ell, j \in [0,k]} [C_{ij} X_{ij} - X_{ij} \log X_{ij}] + \sum_{i \in \ell} [X1]_i, \log [X1]_i \right)$. \hspace{1cm} (5)

For any $q \in \Delta^P_R$, the function $g(X)$ approximates the $\mathbb{P}(q, \phi)$ term and is stated below.

Lemma 3.2 (Theorem 6.7 and Lemma 6.9 in [ACSS21]). Let $R$ be a probability discretization set. For any profile $\phi \in \Phi^n$ with $k$ distinct frequencies the following statements hold for $\alpha = \min(k, |R|) \log n$: $\lim_{n \to \infty} (-O(\alpha)) \cdot C_{\phi} \cdot \max_{X \in Z^n_R} g(X) \leq \max_{q \in \Delta^P_R} \mathbb{P}(q, \phi) \leq \mathbb{P}(\phi, \phi)$.

The proof of concavity for the function $g(X)$ and a running time analysis to solve the convex program are provided in [CSS19a]. For any $X \in Z^n_R$, a pseudo-distributions associated with it is defined below.

Definition 3.3. For any $X \in Z^n_R$, the discrete pseudo-distribution $q_X$ associated with $X$ and $R$ is defined as follows: for arbitrary $[X1]_i$, number of domain elements assign probability $r_i$. Further $p_X \equiv q_X/\|q_X\|_1$ is the distribution associated with $X$ and $R$.

Note that $q_X$ is a valid pseudo-distribution because of the third condition in Equation (3) and these pseudo distributions $p_X$ and $q_X$ satisfy the following lemma.

Lemma 3.4 (Theorem 6.7 in [ACSS21]). Let $R$ and $\phi \in \Phi^n$ be a probability discretization set and a profile with $k$ distinct frequencies. For any $X \in Z^n_R$, the discrete pseudo distribution $q_X$ and distribution $p_X$ associated with $X$ and $R$ satisfy: $\lim_{n \to \infty} (-O(k \log n)) C_{\phi} \cdot g(X) \leq \mathbb{P}(q_X, \phi) \leq \mathbb{P}(p_X, \phi)$.\footnote{The theorem statement in [ACSS21] is only written with an approximation factor of $\exp(O(k \log n))$. However, their proof provides a stronger approximation factor which is upper bounded by the non-negative rank of the probability matrix, which in turn is upper bounded by the minimum of distinct frequencies and distinct probabilities. Therefore the theorem statement in [ACSS21] holds with a much stronger approximation guarantee of $\exp(O(\min(k, |R|) \log n))$.}
4 Approximate PML algorithm

Here we provide a proof sketch of Theorem 2.3 and provide a rounding algorithm that proves it. Our rounding algorithm takes as input a matrix $X \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}$ which may have fractional row sums and round it to integral values. This new rounded matrix $X_{\text{final}}$ corresponds to our approximate PML distribution (See Definition 3.3). The description of our algorithm is as follows.

**Algorithm 1 ApproximatePML($\phi$, $R = \mathbb{R}_{n, \alpha}$)**

1. Let $X$ be any solution that satisfies, $\log g(X) \geq \max_{Y \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}} \log g(Y) - O(\min(k, |R|) \log n)$.
2. $X' = \text{sparsify}(X)$.
3. $(A, B) = \text{swapmatrixround}(X')$.
4. $(X_{\text{final}}, R_{\text{final}}) = \text{create}(A, B, R)$
5. Let $p'$ be the distribution with respect to $X_{\text{final}}$ and $R_{\text{final}}$ (See Definition 3.3).
6. Return $q = \text{discretize}(p', \phi, R)$

We now provide a guarantee for each of these lines of Algorithm 1. We later use these guarantees to prove our final theorem (Theorem 2.3). The guarantees of the approximate maximizer $X$ computed in the first step of the algorithm are summarized in the following lemma.

**Lemma 4.1** (CSS19a, ACSS20). Line 7 of the algorithm can be implemented in $O(|R|k^2 + |R|^2 k)$ time and the approximate maximizer $X$ satisfies: $C_\phi \cdot g(X) \geq \exp(-O(\min(k, |R|) \log n)) \max_{q \in \Delta_R} p(q, \phi)$.

The guarantees of the second step of our algorithm are summarized in the following lemma. Please refer to [ACSS20] for the description of the procedure sparsify. We use this procedure so that we can assume $|R| \leq k+1$ as we can ignore the zero rows of the matrix $X$.

**Lemma 4.2** (Lemma 4.3 in ACSS20). For any $X \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}$, the algorithm $\text{sparsify}(X)$ runs in time $O(|R|k^2)$ and outputs $X' \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}$ such that: $g(X') \geq g(X)$ and $|\{i \in [t] \mid |X'|_i > 0\}| \leq k+1$.

To explain our next step, we need to define a new operation called the swap.

**Definition 4.3.** Given a matrix $A$, indices $i_1 < i_2, j_1 < j_2$ and a parameter $\epsilon \geq 0$, the operation $\text{swap}(A, i_1, i_2, j_1, j_2, \epsilon)$ outputs a matrix $A'$ that satisfies,

$$A'_{ij} = \begin{cases} A_{i,j} + \epsilon & \text{for } i = i_1, j = j_1 \\ A_{i,j} - \epsilon & \text{for } i = i_1, j = j_1 \\ A_{i,j} + \epsilon & \text{for } i = i_2, j = j_2 \\ A_{i,j} - \epsilon & \text{for } i = i_2, j = j_2 \\ A_{i,j} & \text{otherwise.} \end{cases}$$

**Definition 4.4** (Swap distance). $A'$ is $x$-swap distance from $A$, if $A'$ can be obtained from $A$ through a sequence of swap operations and the summation of the value $\epsilon$'s in these operations is at most $x$, i.e. there is a set of parameters $\{(i_1, i_2, j_1, j_2, \epsilon_s)\}_{s \in [t]}$, where $\sum_{s \in [t]} \epsilon_s \leq x$, such that $A' = \text{swap}(A^{(t-1)}, i_1, i_2, j_1, j_2, \epsilon_s)$ for $s \in [t]$, where $A^{(0)} = A$ and $A^{(t)} = A'$.

The following lemma directly follows from Definition 4.3 and Definition 4.4.

**Lemma 4.5.** For any matrices $A, A' \in \mathbb{R}^{s \times t}$, if $A'$ is $x$-swap distance from $A$ for some $x \geq 0$, then $A'^{\top} = A^{\top}$ and $A'^{\top \top} = A^{\top \top}$.

Recall that our objective $g(X)$ contains two terms: (1) the linear term $\sum_{i \in [t], j \in [0, k]} C_{ij} X_{ij}$ and (2) the entropy term $\sum_{i \in [t]} [X^{\top}]_{i} \log[X^{\top}]_{i} - \sum_{i \in [t], j \in [0, k]} X_{ij} \log X_{ij}$. The swap operation always increases the first term, and in the following lemma we bound the loss due to the second term.

**Lemma 4.6.** If $A' \in \mathbb{R}^{s \times [0, k]}$ is $x$-swap distance from $A \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}$, then, $A' \in \mathbb{Z}_R^{\phi, \frac{\text{int}}{\text{frac}}}$ and $g(A') \geq \exp(-O(x \log n)) g(A)$.

One of the main contributions of our work is the following lemma, where we repeatedly apply swap operation to recover a matrix $A$ which exhibits several nice properties as stated below.
Lemma 4.7. For any matrix $A \in \mathbb{R}^{s \times t}$ ($s \leq t$) that satisfies $A^\top 1 = 1$ $\in \mathbb{Z}_{\geq 0}$. The algorithm swapmatrixround runs in $O(s^2t)$ time and returns matrices $A'$ and $B$ such that,

- $A'$ is $O(s)$-swap distance from $A$, $A'^\top 1 = A^\top 1$ and $A'^\top B = A^\top B$.
- $0 \leq B_{ij} \leq A'_{ij}$ for all $i \in [s]$ and $j \in [t]$, $B^\top 1 \in \mathbb{Z}_{\geq 0}^s$, $B'^\top 1 \in \mathbb{Z}_{\geq 0}^s$ and $||A' - B||_1 \leq O(s)$.

The above lemma helps us modify our matrix $X$ to a new matrix $A$ that we can round using the create procedure. The guarantees of this procedure are summarized below.

Lemma 4.8 (Lemma 6.13 in [ACSS21]). For any $A \in \mathbb{R}^{\ell \times [0,k]}$ such that $B \leq A$, $B^\top 1 \in \mathbb{Z}^{\ell}$, $B'^\top 1 \in \mathbb{Z}^{[0,k]}$ and $||A - B||_1 \leq t$. The algorithm create($A, B, R$) runs in time $O(\ell k)$ and returns a solution $A'$ and a probability discretization set $R'$ such that $|R'| \leq |R| + \min(k + 1, t)$, $A' \in \mathbb{R}^{\ell \times [0,k]}$, and $g(A') \geq \exp(-O(t \log n))g(A)$.

As our final goal is to return a distribution in $\Delta_R^D$, we also use the following discretization lemma.

Lemma 4.9. The function discretize takes as input a distribution $p \in \Delta_R$ with $\ell'$ distinct probability values, a profile $\phi$, a discretization set of the form $R = R_{n, \alpha}$ for some $\alpha > 0$ and outputs a pseudo distribution $q \in \Delta_R^D$ such that: $\mathbb{P}(\frac{q}{\|q\|_1}) \geq \exp(-O(\min(k, |R|)) + \min(k, \ell') + \alpha^2 n) \log n)$.

In Section 5, we use the guarantees stated above for each line of Algorithm 1 to prove Theorem 2.3. The description of the function discretize is specified in the proof of Lemma 4.9. We describe the procedure swapmatrixround and provide a proof sketch of Lemma 4.7 in Section 4.1.

4.1 Description of swapmatrixround and comparison to [ACSS20]

Here we describe the procedure swapmatrixround and compare our rounding algorithm to [ACSS20]. Both of [ACSS20] and our approximate PML algorithm have four main lines (1-4); we differ from [ACSS20] in the key Line 4. This line in [ACSS20] invokes a procedure called matrixround that takes as input a matrix $A \in \mathbb{R}^{\ell \times [0,k]}$ and outputs a matrix $B \in \mathbb{R}^{\ell \times [0,k]}$ such that $B \leq A$, $B^\top 1 \in \mathbb{Z}^{\ell}$, $B'^\top 1 \in \mathbb{Z}^{[0,k]}$ and $||A - B||_1 \leq O(\ell + k)$. Such a matrix $B$ is crucial as the procedure create uses $B$ to round fractional row sums of matrix $A$ to integral values. The error incurred in these two steps is at most $\exp(\mathbb{E}[||A - B||_1 \log n]) \in \exp(O((\ell + k) \log n))$. As the procedure sparsifies allows us to assume $\ell \leq k + 1$, we get an $\exp(-k \log n)$ approximate PML using [ACSS20]. However, the setting that we are interested in is when $\ell \ll k$; for instance when $\ell \in O(n^{1/3})$ and $k = O(\sqrt{n})$. In these settings, we desire an $\exp(-O(\min(k, \ell) \log n)) \in \exp(-O(\ell \log n))$ approximate PML. In order to get such an improved approximation using [ACSS20], we need a matrix $B$ satisfying the earlier mentioned inequalities along with $||A - B||_1 \leq O(\min(k, \ell))$. However, such a matrix $B$ may not exist for arbitrary matrices $A$ and the best guarantee any algorithm can achieve is $||A - B||_1 \leq O(\ell + k)$.

To overcome this, we introduce a new procedure called swapmatrixround that takes as input, a matrix $A$ and transforms it to a new matrix $A'$ that satisfies: $g(A') \geq \exp(-O(\min(k, \ell) \log n))g(A)$. Furthermore, this transformed matrix $A'$ exhibits a matrix $B$ that satisfies the guarantees: $B \leq A'$, $B^\top 1 \in \mathbb{Z}_{\geq 0}^\ell$, $B'^\top 1 \in \mathbb{Z}_{\geq 0}^k$ and $||A' - B||_1 \leq O(\ell)$. These matrices $A'$ and $B$ are nice in that we can invoke the procedure create, which would output a valid distribution with required guarantees. In the following we provide a description of the algorithm that finds these matrices $A'$ and $B$.

Algorithm 2 swapmatrixround($A$)

1: Let $A^{(0)} = A$ and $D^{(0)} = 0$.
2: for $r = 1 \ldots \ell$ do
3: \quad $(Y, j) = \text{partialRound}(A^{(r-1)}, r)$
4: \quad $A^{(r)} = \text{roundRow}(Y, j, r)$
5: \quad $D^{(r)} = D^{(r-1)} + Y - A^{(r)}$
6: end for
7: Return $A' = D^{(\ell)} + A^{(\ell)}$ and $B = A^{(\ell)}$. 

8
Our algorithm includes two main subroutines: partialRound and roundiRow. At each iteration $i$, the procedure partialRound considers row $i$ and modifies it by repeatedly applying the swap operation. This modified row is nice as the procedure roundiRow can round this row to have an integral row sum while not affecting the rows in $[i - 1]$. By iterating through all rows, we get the required matrices $A'$ and $B$ that satisfy the required guarantees. In the remainder, we formally state the guarantees achieved by the procedures partialRound and roundiRow.

**Lemma 4.10.** The algorithm partialRound takes as inputs $X \in \mathbb{R}_{\geq 0}^{\times[0,k]}$ and $i \in [\ell-1]$ that satisfies the following, $[X^\top]_{i'} \in \mathbb{Z}_{\geq 0}$ for all $i' \in [1,i-1]$ and $[X^\top]_{j} \in \mathbb{Z}_{\geq 0}$ for all $j \in [0,k]$, and outputs a matrix $Y \in \mathbb{R}_{\geq 0}^{\times[0,k]}$ and an index $j'$ such that:

- $Y$ is within 3-swap distance from $X$.
- $Y_{ij'} \geq o$ and $\sum_{i' = 1}^{i-1} Y_{i'j'} + Y_{ij'} - o \in \mathbb{Z}_{\geq 0}$, where $o = [X^\top]_{i} - [[X^\top]_{i}]$.

Furthermore, the running time of the algorithm is $O(\ell k)$.

Note that by Lemma 4.5 if $Y$ is within 3-swap distance from $X$, then $Y^\top = X^\top$ and $Y^\top = X^\top$.

**Lemma 4.11.** The algorithm roundiRow takes as inputs $Y \in \mathbb{R}_{\geq 0}^{\times[0,k]}$, an column index $j \in [0,k]$ and a row index $i \in [\ell-1]$ such that: $Y^\top X \in \mathbb{Z}_{\geq 0}^{0,k}, Y_{ij} \geq o$ and $\sum_{i' = 1}^{i-1} Y_{i'j} + Y_{ij} - o \in \mathbb{Z}_{\geq 0}$, where $o = [Y^\top]_{i} - [[Y^\top]_{i}]$. Outputs a matrix $X \in \mathbb{R}_{\geq 0}^{\times[0,k]}$ such that,

- $X \leq Y$ and $\|X - Y\|_1 \leq 1$.
- $[X^\top]_{i'} = [Y^\top]_{i'}$ for all $i' \in [i-1]$, $[X^\top]_{j} \in \mathbb{Z}_{\geq 0}$, and $X^\top \in \mathbb{Z}_{\geq 0}^{0,k}$.

We defer the description of all the missing procedures and proofs to appendix.

## 5 Proof of Main Result (Theorem 2.3)

Here we put together the results from the previous sections to prove, Theorem 2.3.

**Proof of Theorem 2.3.** Algorithm 1 achieves the guarantees of Theorem 2.3. In the remainder of the proof, we combine the guarantees of each step of the algorithm to prove the theorem. Toward this end, we first show the following two inequalities: $X_{\text{final}} \in Z_{\mathbb{R},\text{frac}}^{\phi}$, and $g(X_{\text{final}}) \geq \exp(-O(\min(k, |R|) \log n))g(X)$. By Lemma 4.1, the Line 1 of Algorithm 1 returns a solution $X \in Z_{\mathbb{R},\text{frac}}^{\phi}$ that satisfies,

$$C_{\phi} \cdot g(X) \geq \exp(-O(\min(k, |R|) \log n)) \max_{q \in \Delta_{K}^{\phi}} \|p(q, \phi)\|.$$  \hspace{1cm} (7)

By Lemma 4.2, the Line 2 of Algorithm 1 takes input $X$ and outputs $X'$ such that

$$X' \in Z_{\mathbb{R},\text{frac}}^{\phi} \text{ and } g(X') \geq g(X),$$  \hspace{1cm} (8)

and $|\{i \in [\ell] \mid [X'_{j}]_{i} > 0\}| \leq k + 1$. As the matrix $X'$ has at most $k + 1$ non-zero rows, without loss of generality we can assume $|R| \leq k + 1$ (by discarding zero rows).

As matrix $X' \in Z_{\mathbb{R},\text{frac}}^{\phi}$, we have that $X'$ has integral column sums and by invoking Lemma 4.7 with parameters $s = |R|$ and $t = k + 1$, we get matrices $A$ and $B$ that satisfy guarantees of Lemma 4.7. As $A^\top X' = [X'_{j}]$, for all $i \in [\ell]$, $A^\top X = [X^\top X_{j}]$ for all $j \in [0,k]$ and $X' \in Z_{\mathbb{R},\text{frac}}^{\phi}$. We immediately get that $A \in Z_{\mathbb{R},\text{frac}}^{\phi}$. Further note that $A$ is within $O(|R|) = O(\min(|R|, k))$-swap distance from $X'$ and by Lemma 4.6 we get that $g(A) \geq \exp(-O(\min(|R|, k) \log n))g(X')$. To summarize, we showed the following inequalities,

$$A \in Z_{\mathbb{R},\text{frac}}^{\phi} \text{ and } g(A) \geq \exp(-O(\min(|R|, k) \log n))g(X').$$  \hspace{1cm} (9)
Note that, Lemma\textsuperscript{4.7} also outputs a matrix $B$ that satisfies: $B \preceq A$, $B\overset{1}{\rightarrow} \in Z^\prime$, $B\overset{T}{\rightarrow} \in Z^{[0,k]}$ and $\|A - B\|_1 \leq O(\min(|R|,k))$. These matrices $A$ and $B$ satisfy the conditions of Lemma\textsuperscript{4.8} with parameter value $t = O(\min(|R|,k))$. Therefore, the procedure creates takes in input matrices $A$, $B$ and returns a solution $(X_{\text{final}}, R_{\text{final}})$ such that $|R_{\text{final}}| \leq |R| + \min(R,k) \leq 2|R|$ and,

$$X_{\text{final}} \in Z^\prime_{R_{\text{final}}}, \text{ and } g(X_{\text{final}}) \geq \exp(-O(\min(|R|,k) \log n))g(A).$$ \hspace{1cm} (10)

As $X_{\text{final}} \in Z^\prime_{R_{\text{final}}}$, by definition Definition\textsuperscript{3.3} and Lemma\textsuperscript{3.2} the distribution $p^\prime$ satisfies,

$$p(p^\prime, \phi) = \exp(-O(\min(|k|, |R_{\text{final}}|) \log n))C_{\phi}g(X_{\text{final}}) \geq \exp(-O(\min(|k|, |R|) \log n))C_{\phi}g(A) \geq \exp(-O(\min(|k|, |R|) \log n))C_{\phi}g(X) \geq \exp(-O(\min(|k|, |R|) \log n)) \max_{q \in \Delta^\rho_R} \mathbb{P}(q, \phi).$$

In the second inequality we used Equation\textsuperscript{10} and $|R_{\text{final}}| \leq 2|R|$. In the third, fourth and fifth inequalities, we used Equation\textsuperscript{9}, Equation\textsuperscript{8} and Equation\textsuperscript{7} respectively.

Recall we need a distribution that approximately maximizes $\max_{q \in \Delta^\rho_R} \mathbb{P}(\frac{q}{||q||_1}, \phi)$ instead of just $\max_{q \in \Delta^\rho_R} \mathbb{P}(q, \phi)$. In the remainder of the proof we provide a procedure to output such a distribution.

For any constant $c > 0$, let $c \cdot R \overset{\text{def}}{=} \{c \cdot r_i | r_i \in R\}$. For any $q \in \Delta^\rho_R$, as $||q||_1$ satisfies: $r_{\text{min}} \leq ||q||_1 \leq c$, we get that,

$$\max_{q \in \Delta^\rho_R} \mathbb{P}\left(\frac{q}{||q||_1}, \phi\right) = \max_{c \in [1,1/r_{\text{min}}]} \max_{q \in \Delta^\rho_R} \mathbb{P}(q, \phi).$$ \hspace{1cm} (11)

The above expression holds as the maximizer $q^*$ of the left hand side satisfies: $q^* \in \Delta^\rho_R(\{1 + \beta\}^\prime)$, where $\beta \in o(1)$, where $a = O(\frac{1}{\log(1/r_{\text{min}})})$ is such that $r_{\text{min}}(1 + \beta)^n = 1$. For any constant $c \in [1,1/r_{\text{min}}]$, note that there exists a constant $c' \in C$ such that $c(1-\beta) \leq c' \leq c$. Furthermore, for any distribution $q_0 \in \Delta^\rho_R$ with $||q_0||_1 = 1/c$, note that the distribution $q' = c'q \in \Delta^\rho_R$ and satisfies: $\mathbb{P}(\frac{q}{||q||_1}, \phi) = \mathbb{P}(c \cdot q, \phi) = \mathbb{P}(\frac{q'}{||q'||_1}, \phi) = (\frac{c'}{c})^n \mathbb{P}(q', \phi)$.

Therefore we get that, $\mathbb{P}(q', \phi) = (\frac{c'}{c})^n \mathbb{P}(\frac{q'}{||q'||_1}, \phi) \geq (1-\beta)^n \mathbb{P}(\frac{q}{||q||_1}, \phi) \geq \exp(-2\beta n) \mathbb{P}(\frac{q}{||q||_1}, \phi)$.

Combining this analysis with Equation\textsuperscript{11} we get that,

$$\max_{c \in C} \max_{q \in \Delta^\rho_R} \mathbb{P}(q, \phi) \geq \exp(-2\beta n) \max_{q \in \Delta^\rho_R} \mathbb{P}(\frac{q}{||q||_1}, \phi).$$ \hspace{1cm} (12)

For each $c > 0$, $|R| = c \cdot R$, our algorithm (Algorithm\textsuperscript{1}) returns a distribution $p_c$ that satisfies,

$$\mathbb{P}(p_c, \phi) \geq \exp(-O(\min(k, |R|) \log n)) \max_{q \in \Delta^\rho_R} \mathbb{P}(q, \phi).$$

Let $p^*$ be the distribution that achieves the maximum objective value to our convex program among the distributions $\{p_c\} \in C$. Then note that $p^*$ satisfies: $\mathbb{P}(p^*, \phi) \geq \exp(-O(\min(k, |R|) \log n)) - 2\beta n \max_{q \in \Delta^\rho_R} \mathbb{P}(\frac{q}{||q||_1}, \phi)$. Substituting $\beta = \frac{\min(k, |R|)}{n}$ in the previous expression, we get,

$$\mathbb{P}(p^*, \phi) \geq \exp(-O(\min(k, |R|) \log n)) \max_{q \in \Delta^\rho_R} \mathbb{P}(\frac{q}{||q||_1}, \phi).$$

As each of our distributions $p_c$ (including $p^*$) have the number of distinct probability values upper bounded by $2|R|$, by applying Lemma\textsuperscript{4.9} we get a pseudo distribution $q \in \Delta^\rho_R$ with the desired guarantees. The final run time of our algorithm is $O(\frac{|C|}{10}) \in O(\frac{n}{\min(k, |R|)} \cdot T_{\text{final}})$, where $T_{\text{final}}$ is the time to implement Algorithm\textsuperscript{1}. Further note that by Lemma\textsuperscript{3.1} without loss of generality we can assume $|R| \leq n/k$. As all the lines of Algorithm\textsuperscript{1} are polynomial in $n$, our final running time follows from the run times of each line and we conclude the proof.

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References


Checklist

1. For all authors...
   (a) Do the main claims made in the abstract and introduction accurately reflect the paper’s contributions and scope? [Yes]
   (b) Did you describe the limitations of your work? [Yes] See Section 2
   (c) Did you discuss any potential negative societal impacts of your work? [N/A]
   (d) Have you read the ethics review guidelines and ensured that your paper conforms to them? [Yes]

2. If you are including theoretical results...
   (a) Did you state the full set of assumptions of all theoretical results? [Yes]
   (b) Did you include complete proofs of all theoretical results? [Yes]

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   (a) Did you include the code, data, and instructions needed to reproduce the main experimental results (either in the supplemental material or as a URL)? [N/A]
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A Additional definitions

Optimal sample complexity The sample complexity of an estimator \( \hat{f} : \mathcal{D}^n \to \mathbb{R} \) when estimating a property \( f : \Delta^P \to \mathbb{R} \) for distributions in a collection \( P \subseteq \Delta^P \), is the number of samples \( \hat{f} \) needs to determine \( f \) with accuracy \( \epsilon \) and low failure probability \( \delta \) for all distributions in \( P \). Specifically,

\[
C^\hat{f}(f, P, \delta, \epsilon) \overset{\text{def}}{=} \min\{n \mid \mathbb{P}\left( |f(p) - \hat{f}(a^n)| \geq \epsilon \right) \leq \delta \text{ for all } p \in P\}.
\]

The sample complexity of estimating \( f \) is the lowest sample complexity of any estimator,

\[
C^*(f, P, \delta, \epsilon) = \min_{\hat{f}} C^\hat{f}(f, P, \delta, \epsilon).
\]

In the paper, the dependency on \( \delta \) is typically de-emphasized and \( \delta \) is assumed to be 1/3.

B Proof of Theorem 2.2

Here we provide the proof of Theorem 2.2. The proof of this statement is implicit in the analysis presented in [HS21]. Here we provide a simpler and short proof that uses the continuity lemma presented in [HS21]. For convenience, we restate this key lemma in our notation.

**Lemma B.1 (Lemma 2 in [HS21])** Let \( A \geq 2, c_0 \in (0, 1), r, s \) be arbitrary constants with \( 0 < s < r < 1/2 \) and let \( R = R_{c_0, r} \). Then there exists a constant \( c = c(A, c_0, r, s) > 0 \) such that for any distribution \( p \in \Delta^P \), there exists a pseudo distribution \( q \in \Delta^P_R \) such that: for all \( S \subseteq \Phi^n \), it holds that,

\[
\mathbb{P}(p, S) \geq \mathbb{P}\left(\frac{q}{\|q\|_1}, S\right)^{1/(1-c_0 n^{-s})} \exp(-cn^{-2r+s})
\]

\[
\mathbb{P}\left(\frac{q}{\|q\|_1}, S\right) \geq \mathbb{P}(p, S)^{1/(1-c_0 n^{-s})} \exp(-cn^{-2r+s})
\]

**Proof of Theorem 2.2** Let \( p \) be the hidden distribution. Given a profile \( \phi \), let \( q_\phi \in \Delta^P_R \) be any pseudo distribution that satisfies,

\[
\mathbb{P}\left(\frac{q_\phi}{\|q_\phi\|_1}, \phi\right) \geq \exp(-O(|R| \log n)) \max_{q \in \Delta^P_R} \mathbb{P}\left(\frac{q}{\|q\|_1}, \phi\right)
\]

\[
\geq \exp(-O(n^{-1/3} \log^2 n)) \max_{q \in \Delta^P_R} \mathbb{P}\left(\frac{q}{\|q\|_1}, \phi\right).
\]

As \( L \overset{\text{def}}{=} |\Delta^P_R| \leq \exp(O(n^{1/3} \log^2 n)) \), we use \( q_1, \ldots, q_L \) to denote the pseudo distributions in \( \Delta^P_R \).

Let \( G = \{ \phi \in \Phi^n \mid |f(p) - \hat{f}(\phi)| < \epsilon \} \), that is, the set of all profiles where the estimator succeeds.

Also let \( S_i = \{ \phi \in G \mid q_\phi = q_i \} \). Using these definitions in the remainder of the proof, we upper bound the failure probability of our estimator.

\[
\mathbb{P}\left(\left| f\left(\frac{q_\phi}{\|q_\phi\|_1}\right) - f(p) \right| > \epsilon \right) \leq \mathbb{P}(p, \Phi^n \setminus G) + \sum_{\{i \in [1:L] \mid |f\left(\frac{q_i}{\|q_i\|_1}\right) - f(p)| > \epsilon\}} \mathbb{P}(p, S_i),
\]

\[
\leq \delta + \sum_{\{i \in [1:L] \mid |f\left(\frac{q_i}{\|q_i\|_1}\right) - f(p)| > \epsilon\}} \mathbb{P}(p, S_i).
\]

In the above inequality we used that the failure probability of the estimator is at most \( \delta \), that is, \( \mathbb{P}(p, \Phi^n \setminus G) \leq \delta \). First note that from the definitions of \( q_i \), \( S_i \) and \( q_\phi \), we have that,

\[
\mathbb{P}\left(\frac{q_i}{\|q_i\|_1}, S_i\right) \geq \exp(-O(n^{-1/3} \log^2 n)) \max_{q \in \Delta^P_R} \mathbb{P}\left(\frac{q}{\|q\|_1}, S_i\right).
\]

Further applying Lemma B.1 with \( r = 1/3 \) and \( s \) with any tiny constant, we get that,

\[
\mathbb{P}\left(\frac{q_i}{\|q_i\|_1}, S_i\right) \geq \exp(-O(n^{-1/3} \log n)) \mathbb{P}(p, S_i)^{1+O(n^{-s'})} \exp(-O(n^{1/3+c'}))
\]
for any tiny constant \( c' > 0 \). The above expression further simplifies to,

\[
P\left( \frac{q_i}{\|q_i\|_1}, S_i \right) \geq \exp(-O(n^{1/3+c'}))P\left( p, S_i \right)^{1+O(n^{-c'})}.
\]

Suppose \( S_i \) is set such that, \( P(p, S_i) \geq \delta \), then note that \( P\left( \frac{q_i}{\|q_i\|_1}, S_i \right) > \delta \). As the estimator provided by the conditions of the lemma succeeds on all the profiles in \( S_i \), we that \( |f(p) - \tilde{f}(\phi)| \leq \epsilon \) for all \( \phi \in S_i \). Suppose \( |f(p) - \tilde{f}(\phi)| > 2\epsilon \), then by triangle inequality this would imply \( |f(p) - \tilde{f}(\phi)| > \epsilon \) for all \( \phi \in S_i \). However note that, \( P\left( \frac{q_i}{\|q_i\|_1}, S_i \right) > \delta \), and this would imply that the failure probability of the estimator is greater than \( \delta \) when the underlying distribution is \( \frac{q_i}{\|q_i\|_1} \); a contradiction. Therefore, it should be the case that \( |f(p) - \tilde{f}(\phi)| \leq 2\epsilon \) for all \( S_i \) that satisfy \( P(p, S_i) \geq \delta \); and our failure probability is upper bounded by,

\[
P\left( |f\left( \frac{q_i}{\|q_i\|_1}\right) - f(p)| > \epsilon \right) \leq \delta + |L|\delta^{1-O(n^{-c'})} \exp(O(n^{1/3+c'})).
\]

Further, substituting the value of \(|L| \leq \exp(O(n^{1/3} \log^2 n))\), we get our desired result and we conclude our proof. \( \square \)

**C Other results**

**Lemma C.1.** For any two vectors \( u, v \in \mathbb{R}^{[0,k]} \), the following inequality holds,

\[
(u^T \mathbf{1} \log u^T \mathbf{1} + v^T \mathbf{1} \log v^T \mathbf{1}) - \sum_{i \in [0,k]} (u_i \log u_i + v_i \log v_i) \leq w^T \mathbf{1} \log w^T \mathbf{1} - \sum_{i \in [0,k]} w_i \log w_i
\]

where \( w = u + v \).

**Proof.** For any \( x \in \mathbb{R}^{[0,k]} \), let \( f(x) \overset{\text{def}}{=} x^T \mathbf{1} \log x^T \mathbf{1} - \sum_{i \in [0,k]} x_i \log x_i \). Note that \( f(x) \) is concave [CSS19a] and furthermore, \( f(c \cdot x) = c \cdot f(x) \).

Let \( w' = \frac{1}{2}u + \frac{1}{2}v \), applying concavity we get that,

\[
f(w') \geq \frac{1}{2}f(u) + \frac{1}{2}f(v),
\]

As \( f(w) = 2f(w') \), combined with above inequality we have our proof. \( \square \)

**Lemma C.2.** For many matrices \( X, Y \in \mathbb{R}^{\ell \times [0,k]} \) (where \( \|X\|_1, \|Y\|_1, k, \ell \leq O(n^2) \)) such that \( \|X - Y\|_1 \leq \alpha \), we have that,

\[
\left| \sum_{i \in \ell, j \in [0,k]} X_{ij} \log X_{ij} - \sum_{i \in \ell, j \in [0,k]} Y_{ij} \log Y_{ij} \right| \leq O(\alpha \log n) + O(\log n).
\]

Furthermore,

\[
\left| \sum_{i \in \ell} [X^T]_i \log [X^T]_i - \sum_{i \in \ell} [Y^T]_i \log [Y^T]_i \right| \leq O(\alpha \log n) + O(\log n).
\]

**Proof.** Note that \( x \log x \) is \( O(\log n) \)-Lipschitz for \( x \geq \frac{1}{n^2} \), and for any \( 0 \leq x_1, x_2 \leq \frac{1}{n^2} \), we have \( |x_1 \log x_1 - x_2 \log x_2| \leq O(\frac{1}{n^2} \log n) \). As a result, for any non-negative \( x_1, x_2 \) that are \( O(n^2) \), we have

\[
|x_1 \log x_1 - x_2 \log x_2| \leq \left( \frac{1}{n^2} + |x_1 - x_2| \right) \cdot O(\log n).
\]

16
Note that \( \sum_{i \in \ell, j \in [0, k]} X_{ij} \leq O(n^2) \) and \( \sum_{i \in \ell, j \in [0, k]} Y_{ij} \leq O(n^2) \). We have
\[
\left| \sum_{i \in \ell, j \in [0, k]} X_{ij} \log X_{ij} - \sum_{i \in \ell, j \in [0, k]} Y_{ij} \log Y_{ij} \right| \leq \sum_{i \in \ell, j \in [0, k]} |X_{ij} \log X_{ij} - Y_{ij} \log Y_{ij}|
\]
\[
\leq \sum_{i \in \ell, j \in [0, k]} \left( \frac{1}{n^2} + |X_{ij} - Y_{ij}| \right) O(\log n) \leq (1 + \|X - Y\|_1) \cdot O(\log n) \leq O(\alpha \log n) + O(\log n)
\]
and similarly,
\[
\left| \sum_{i \in \ell} [X^T]_i \log [X^T]_i, - \sum_{i \in \ell} [Y^T]_i \log [Y^T]_i \right| \leq \sum_{i \in \ell} \left( \frac{1}{n^2} + ||X^T - Y^T||_1 \right) O(\log n) \leq (1 + ||X - Y||_1) O(\log n) \leq O(\alpha \log n) + O(\log n)
\]
We conclude the proof. \( \square \)

**D Description and guarantees of partialRound**

Here we provide the description of partialRound and its guarantees are summarized in Lemma 4.10.

**Algorithm 3** partialRound(X, i)

1: \( Z = \text{split}(X, i) \).
2: \( (W, j) = \text{partial\_round\_special}(Z, i + 1) \).
3: \( Y = \text{combine}(W, i) \).
4: Return \( (Y, j) \).

**Lemma 4.10.** The algorithm partialRound takes as inputs \( X \in \mathbb{R}^{\ell \times [0, k]} \) and \( i \in [\ell - 1] \) that satisfies the following, \( [X^T]_{i'} \in \mathbb{Z}_{\geq 0} \) for all \( i' \in [1, i - 1] \) and \( [X^T Y]_j \in \mathbb{Z}_{\geq 0} \) for all \( j \in [0, k] \), and outputs a matrix \( Y \in \mathbb{R}^{\ell \times [0, k]} \) and an index \( j' \) such that:

- \( Y \) is within 3-swap distance from \( X \).
- \( Y_{i, j'} \geq o \) and \( \sum_{i' = 1}^{i - 1} Y_{i', j'} + Y_{i, j'} - o \in \mathbb{Z}_{\geq 0} \), where \( o = [X^T]_i - [[X^T]_i] \).

Furthermore, the running time of the algorithm is \( O(\ell k) \).

Our algorithm invokes several subroutines and in the following we provide description for each of these and also state their guarantees. The subroutine partial_round_special provides an algorithm that proves the Lemma 4.10 in the special case of \( [X^T]_i \in [0, 1] \). The description of this subroutine and its guarantees are stated below. For convenience, we first introduce some new notation and a new operation trans.

**Notation:** For any matrix \( X \in \mathbb{R}^{\ell \times [0, k]} \) and indices \( i_1, i_2, j_1, j_2 \), we define \( X(i_1 : i_2, j_1 : j_2) \) \( \overset{\text{def}}{=} \sum_{i=i_1}^{i_2} \sum_{j=j_1}^{j_2} X_{ij} \) and say \( (i, j) \in [i_1 : i_2, j_1 : j_2] \) if \( i_1 \leq i \leq i_2 \) and \( j_1 \leq j \leq j_2 \). Therefore, \( X(i_1 : i_2, j_1 : j_2) = \sum\{ (i, j) \in [i_1 : i_2, j_1 : j_2] \} X_{ij} \). Also note that when \( i_2 = i_1 + 1 \) or \( j_2 = j_1 + 1 \), we are summing over a empty set and \( X(i_1 : i_2, j_1 : j_2) \) \( \overset{\text{def}}{=} 0 \).

**Lemma D.1 (Guarantees of trans).** Let \( X \in \mathbb{R}^{\ell \times [0, k]} \). For any indices \( i_1 \leq i_2 + 1 \leq i_3 \leq i_4 + 1 \), \( j_1 \leq j_2 + 1 \leq j_3 \leq j_4 + 1 \), and \( 0 \leq v \leq \min\{X(i_3 : i_4, j_3 : j_4), X(i_1 : i_2, j_3 : j_4)\} \), there exists \( Y \in \mathbb{R}^{\ell \times [0, k]} \) such that, \( Y \) is within \( v \)-swap distance from \( X \), and

- \( Y(i_1 : i_2, j_1 : j_2) = X(i_1 : i_2, j_1 : j_2) + v \)
- \( Y(i_3 : i_4, j_3 : j_4) = X(i_3 : i_4, j_3 : j_4) + v \)
- \( Y(i_1 : i_2, j_3 : j_4) = X(i_1 : i_2, j_3 : j_4) - v \)
- \( Y(i_3 : i_4, j_1 : j_2) = X(i_3 : i_4, j_1 : j_2) - v \)
- \( Y_{ij} = X_{ij} \forall (i, j) \notin [i_1 : i_2, j_1 : j_2] \cup [i_1 : i_2, j_3 : j_4] \cup [i_3 : i_4, j_1 : j_2] \cup [i_3 : i_4, j_3 : j_4] \).
Furthermore, we define operation \( \text{trans}(X, v, i_1, i_2, i_3, i_4, j_1, j_2, j_3, j_4) \) whose output is \( Y \) that satisfies properties above, and the running time of process \( \text{trans} \) is \( O(\ell k) \).

We defer the proof for lemma D.1 to appendix D.1. The description of our subroutine \text{partial\_round\_special} is as follows.

**Algorithm 4** \text{partial\_round\_special}(X, i)

1: \( j = \min\{r : X(1 : \ell, 0 : r) > X(1 : i - 1, 0 : k)\} \).
2: \( t = X(1 : i - 1, 0 : k) - X(1 : \ell, 0 : j - 1) \).
3: \( Y = X \).
4: \( Y = \text{trans}(Y, Y(i : i, 0 : j - 1), 1, i - 1, i, i, 0, j - 1, j, k) \).
5: \text{if } Y(1 : i - 1, j : j) \geq t \text{ then}
6: \( v = Y(1 : i - 1, j : j) - [Y(1 : i - 1, j : j)] \).
7: \( Y = \text{trans}(Y, \min \{v + 1, Y(1 : i - 1, j : j) - t\}, 1, i - 1, i + 1, \ell, 0, j - 1, j, j) \).
8: \text{else}
9: \( v = [Y(1 : i - 1, j : j)] - Y(1 : i - 1, j : j) \).
10: \( Y = \text{trans}(Y, v, 1, i - 1, i, \ell, j, j, j + 1, k) \).
11: \text{end if}
12: \( Y = \text{trans}(Y, Y(i : i, j + 1 : k), i, i, i + 1, \ell, j, j, j + 1, k) \).
13: \text{Return } (Y, j) .

**Lemma D.2** (Guarantee of partial\_round\_special). The algorithm \text{partial\_round\_special} takes in inputs \( X \in \mathbb{R}_{\geq 0}^{l \times [0,k]} \) and \( i \in [1, \ell - 1] \) that satisfy the conditions of Lemma 4.10 with an additions assumption that \( \sum_{j' \in [0,k]} X_{ij'} \in [0, 1] \) and outputs a matrix \( Y \in \mathbb{R}_{\geq 0}^{l \times [0,k]} \) and an index \( j \) that satisfy the guarantees of Lemma 4.10. Furthermore, the running time of this procedure is \( O(\ell k) \).

We defer the proof for lemma D.2 to appendix D.2. To extend the above algorithm for the general case, we define simple operations split and combine that we define next. Intuitively, the operation combine combines \( i \)-th and \((i + 1)\)-th rows by adding them up, and split splits \( i \)-th row into two rows so that the summation of one row is an integer and the summation of the other row is less than one.

**Definition D.3** (Combine). For any \( W \in \mathbb{R}_{\geq 0}^{(l + 1) \times [0,k]} \) and \( t \in [1, \ell] \), \( Y \overset{\text{def}}{=} \text{combine}(W, t) \in \mathbb{R}_{\geq 0}^{l \times [0,k]} \) is defined as follows,

\[
Y_{i,j} = \begin{cases} 
W_{i,j} & i \leq t - 1 \\
W_{i,j} + W_{(i+1),j} & i = t \\
W_{(i+1),j} & i \geq t + 1
\end{cases}
\]

![Figure 1: trans operation. It increases two blocks by \( v \) respectively, and decreases two blocks by \( v \) respectively.](image-url)
We use a $3 \times 3$ partitioned matrix to denote $Y \in \mathbb{R}^\ell \times [0, k]$, where the upper (middle / lower) row denotes the first $(i - 1)$ rows (the $i$-th row / the last $(\ell - i)$ rows) of $Y$, and the left (middle / right) column denotes the 0-th to $(j - 1)$-th columns (the $j$-th column / the last $(k - j)$ columns) of $Y$. Symbol $+$ or $-$ means increasing or decreasing some elements in the block. Symbol $= 0$ means all the elements in the block are zero. Symbol $\in \mathbb{Z} \geq 0$ means the sum of the block is a non-negative integer.

**Figure 2:** We use a $3 \times 3$ partitioned matrix to denote $Y \in \mathbb{R}^\ell \times [0, k]$, where the upper (middle / lower) row denotes the first $(i - 1)$ rows (the $i$-th row / the last $(\ell - i)$ rows) of $Y$, and the left (middle / right) column denotes the 0-th to $(j - 1)$-th columns (the $j$-th column / the last $(k - j)$ columns) of $Y$. Symbol $+$ or $-$ means increasing or decreasing some elements in the block. Symbol $= 0$ means all the elements in the block are zero. Symbol $\in \mathbb{Z} \geq 0$ means the sum of the block is a non-negative integer.

**Figure 3:** combine operation. It combines the $t$-th row and the $(t + 1)$-th row by adding them up

**Definition D.4 (Split).** Let $X \in \mathbb{R}^\ell \times [0, k]$, $r(t) := \sum_{j \in [0, k]} x_{tj}$ and $s(t) := \min\{j : \sum_{j' \leq j} x_{tj'} \geq |r(t)|\}$. We define split($X, t$) to be $Z \in \mathbb{R}^{(\ell + 1) \times [0, k]}$ where,

$$Z_{i, j} = X_{i, j}, \quad i \leq t - 1$$

$$Z_{i, j} = X_{(i - 1), j}, \quad i \geq t + 2$$

\[
\begin{align*}
Z_{t, j} &= \begin{cases} 
X_{t, j}, & j < s(t) \\
|r(t)| - \sum_{j' \leq j-1} X_{t, j'}, & j = s(t) \\
0, & j > s(t)
\end{cases} \\
Z_{t+1, j} &= \begin{cases} 
0, & j < s(t) \\
\sum_{j' \leq j} X_{t, j'} - |r(t)|, & j = s(t) \\
X_{t, j}, & j > s(t)
\end{cases}
\end{align*}
\]

We remark that $Z$ has the following properties: 1) combine($Z, t$) = $X$, 2) $\sum_{j \in [0, k]} Z_{tj} = |r(t)|$, 3) $\sum_{j \in [0, k]} Z_{(t+1), j} = r(t) - |r(t)|$, which is a real number in $[0, 1)$. 

19
When is a feasible output because of the following properties.

Now we assume the lemma is true when \( s \). We define \( r(X) \) to be \( \{(i, j) \in [i_3 : i_4, j_1 : j_2] \mid X_{ij} > 0\} \), and \( s(X) \) to be \( \{(i, j) \in [i_1 : i_2, j_3 : j_4] \mid X_{ij} > 0\} \). Note that \( r(X) \) is also dependent on \( i_3, i_4, j_1 \) and \( j_2 \), but they are fixed in this proof, and \( s(X) \) is also dependent on \( i_1, i_2, j_3 \) and \( j_4 \). We prove the lemma by induction.

When \( |r(X)| + |s(X)| \leq 1 \), one of \( |r(X)| \) and \( |s(X)| \) is 0, so \( v = 0 \), and \( Y = X \) is a feasible output.

Now we assume the lemma is true when \( |r(X)| + |s(X)| < t \), and show it is still true when \( |r(X)| + |s(X)| = t \), where \( t \geq 2 \).

If \(|r(X)| = 0\) or \(|s(X)| = 0\), we have \( v = 0 \), and \( Y = X \) is a feasible output.

If \(|r(X)| > 0\) and \(|s(X)| > 0\), we draw arbitrary elements \((c, b) \in r(X)\) and \((a, d) \in s(X)\).

Let \( u = \min \{X_{cb}, X_{ad}, v\} \). Let \( W = \text{swap}(X, a, c, b, d, u) \), and we have

\[
\begin{align*}
W &= u\text{-swap distance from } X, \\
W(i_1 : i_2, j_1 : j_2) &= X(i_1 : i_2, j_1 : j_2) + u \\
W(i_3 : i_4, j_3 : j_4) &= X(i_3 : i_4, j_3 : j_4) + u,
\end{align*}
\]

\[
\begin{align*}
W(i_1 : i_2, j_3 : j_4) &= X(i_1 : i_2, j_3 : j_4) - u \\
W(i_3 : i_4, j_1 : j_2) &= X(i_3 : i_4, j_1 : j_2) - u.
\end{align*}
\]

\[
\begin{align*}
\forall (i, j) \notin [i_1 : i_2, j_1 : j_2] &\cup [i_1 : i_2, j_3 : j_4] & \cup [i_3 : i_4, j_1 : j_2] &\cup [i_3 : i_4, j_3 : j_4], W_{ij} = X_{ij}.
\end{align*}
\]

If \( u = v \), \( Y = W \) is a feasible output. Otherwise, either \( u = X_{cb} \) or \( u = X_{ad} \), so either \( W_{cb} \) or \( W_{ad} \) becomes zero. Note that \((a, b)\) and \((c, d)\) are not in \([i_1 : i_2, j_3 : j_4] \cup [i_3 : i_4, j_1 : j_2] \), so we have \(|r(W)| + |s(W)| \leq |r(X)| + |s(X)| - 1 \). Let \( Y = \text{trans}(W, v - u, i_1, i_2, i_3, i_4, j_1, j_2, j_3, j_4) \), and it is a feasible output because of the following properties.

\[
\begin{align*}
Y &= (u + (v - u) = v)\text{-swap distance from } X, \\
Y(i_1 : i_2, j_1 : j_2) &= W(i_1 : i_2, j_1 : j_2) + (v - u) = X(i_1 : i_2, j_1 : j_2) + v, \\
Y(i_3 : i_4, j_3 : j_4) &= W(i_3 : i_4, j_3 : j_4) + (v - u) = X(i_3 : i_4, j_3 : j_4) + v, \\
Y(i_1 : i_2, j_3 : j_4) &= W(i_1 : i_2, j_3 : j_4) - (v - u) = X(i_1 : i_2, j_3 : j_4) - v, \\
Y(i_3 : i_4, j_1 : j_2) &= W(i_3 : i_4, j_1 : j_2) - (v - u) = X(i_3 : i_4, j_1 : j_2) - v, \\
\forall (i, j) \notin [i_1 : i_2, j_1 : j_2] &\cup [i_1 : i_2, j_3 : j_4] & \cup [i_3 : i_4, j_1 : j_2] &\cup [i_3 : i_4, j_3 : j_4], Y_{ij} = W_{ij} = X_{ij}.
\end{align*}
\]
where the relation between $Y$ and $W$ follows from inductive assumption, and the relation between $W$ and $X$ follows from Equation (13) to (16).

To implement $\text{trans}$, it is sufficient to enumerate all the elements in $r(X)$ and $s(X)$, so the running time can be $O(\ell k)$.

\[
\square
\]

### D.2 Proof for Lemma D.2

Lemma D.2 is a special case of Lemma 4.10 with an additional input constraint $\sum_{j' \in [0, k]} X_{ij'} \in [0, 1]$. In the output $Y$, the $i$-th row only has one positive element $Y_{ij}$. In Algorithm 4, Line 4 makes $Y_{ij'}$ where $j' < j$ be zero, Line 12 makes $Y_{ij'}$ where $j' > j$ be zero, and Line 7 and Line 10 make $\sum_{i' = 1}^{i-1} Y_{i'j}$ to be an integer.

**Proof for Lemma D.2** There are mainly two parts in this proof. The former part proves the properties of output $Y$ based on Lemma D.1, and the latter part proves the parameters are valid when calling $\text{trans}$ in Algorithm 4.

**First part of the proof** In $\text{partial\_round\_special}(X, i)$, $Y(i : i, 0 : j - 1)$ only changes at Line 4 when it decreases by $Y(i : i, 0 : j - 1) = 0$ at Line 13. $Y(i : i, j + 1 : k)$ does not change after Line 12 where it decreases by $Y(i : i, j + 1 : k)$, so we have $Y(i : i, j + 1 : k) = 0$ at Line 13. Note that $Y_{ij} = \sum_{j' \in [0, k]} X_{ij'}$, so the output $Y$ satisfies

\[
Y_{ij} = \sum_{j' \in [0, k]} X_{ij'}
\]

$Y(1 : i - 1, j : j)$ does not change after Line 7 or Line 10. If the algorithm goes to Line 7, $Y(1 : i - 1, j : j)$ decreases by $v + 1$ or $Y(1 : i - 1, j : j) - t$, which means $Y(1 : i - 1, j : j) = \max \{ (\lfloor Y(1 : i - 1, j : j) \rfloor - 1) + t \}$ at Line 13. If the algorithm goes to Line 10, $Y(1 : i - 1, j : j)$ increases by $v$, which means $Y(1 : i - 1, j : j) = \lfloor Y(1 : i - 1, j : j) \rfloor$ at Line 13. In both cases, $Y(1 : i - 1, j : j)$ is a non-negative integer at Line 13, so the output $Y$ satisfies

\[
\sum_{i' = 1}^{i-1} Y_{i'j} \in \mathbb{Z}_{\geq 0}
\]

We use three $\text{trans}$ operations on $Y$ which initially equals to $X$, and we have $v \leq 1$ in each operation $\text{trans}(X, v, i_1, i_2, k_1, k_2)$, so $Y$ is 3-swap distance from $X$.

We use $O(1)$ $\text{trans}$ operations on $Y$, so the algorithm takes $O(\ell k)$ time.

**Second part of the proof** We also need to show $0 \leq v \leq \min\{X(i_3 : i_4, j_1 : j_2), X(i_3 : i_2, j_3 : j_4)\}$ in each operation $\text{trans}(X, v, i_1, i_2, k_1, k_2)$ to ensure the operations are valid.

Based on the definition of $j$ and $t$, we have

\[
X(1 : i - 1, 0 : k) = X(1 : \ell, 0 : j - 1) + t,
0 \leq t < Y(1 : \ell, j : j).
\]

Note that $t$ and $Y(1 : \ell, j : j)$ are both integers, so we have

\[
t \leq Y(1 : \ell, j : j) - 1.
\]

Before Line 4, $Y = X$, and we have $Y(1 : i - 1, 0 : k) \geq Y(1 : \ell, 1 : j - 1)$, which implies

\[
Y(1 : i - 1, j : k) \geq Y(1 : i, 0 : j - 1) \geq Y(i : i, 0 : j - 1),
\]

which implies Line 4 is valid.

Before Line 7, we have $Y(1 : i - 1, 0 : k) = Y(1 : \ell, 0 : j - 1) + t$. Also, we have $Y(i : i, 0 : j - 1) = 0$ because of Line 4. So, we have

\[
Y(i + 1 : \ell, 0 : j - 1) \geq Y(1 : i - 1, j : j) - t.
\]
which implies Line \[7\] is valid.

Before Line \[10\] we have \(Y(1: \ell, j : j) \geq t\). Note that \(t\) is an integer, so when \(Y(1: i - 1, j : j) < t\), we have

\[
Y(i : \ell, j : j) = Y(1: \ell, j : j) - Y(1: i - 1, j : j) \geq t - Y(1: i - 1, j : j) \geq v.
\]

We also have

\[
Y(1: i - 1, j + 1 : k) = Y(1: i - 1, 0 : k) - Y(1: i - 1, 0 : j - 1) - Y(1: i - 1, j : j) \\
= (Y(1: i - 1, 0 : k) - Y(1: \ell, 0 : j - 1)) - Y(1: i - 1, j : j) \\
= t - Y(1: i - 1, j : j) \geq v.
\]

These two formulas imply Line \[10\] is valid.

After Line \[7\] we have either \(Y(1: i - 1, j : j) = t\) (when \(v + 1 < Y(1: i - 1, j : j) - t\)) or \(Y(i + 1 : \ell, j : j) \geq 1\) (the other case). After Line \[10\] we have \(Y(1: i - 1, j : j) \leq t\). Since \(t\) is an integer, when \(Y(1: i - 1, j : j) \leq t\), we have

\[
Y(i : \ell, j : j) = Y(1: \ell, j : j) - Y(1: i - 1, j : j) \geq 1.
\]

As a result, we always have \(Y(i : \ell, j : j) \geq 1\) before Line \[12\]. Note that \(Y(i : i, 0 : k) < 1\) and \(Y(i : i, 0 : j - 1) = 0\), so we have

\[
Y(i + 1 : \ell, j : j) = Y(i : \ell, j : j) - Y_{ij} > Y(i : i, 0 : k) - Y_{ij} = Y(i : i, j + 1 : k),
\]

which implies Line \[12\] is valid.

\[\square\]

### D.3 Proof for Lemma 4.10

**Proof of Lemma 4.10** We prove the properties mentioned in Lemma 4.10 one by one. Firstly we have

- \(Y_{ij} \geq W_{(i+1),j} = \sum_{j' \in [0,k]} Z_{(i+1),j'} = o,\)

- \(\sum_{i' = 1}^{i-1} Y_{i'j} + Y_{ij} - o = \sum_{i' = 1}^{i+1} W_{i'j} - o = \sum_{i' = 1}^{i} W_{i'j} \in \mathbb{Z}_{\geq 0}.

where the relation between \(W\) and \(Z\) is from Lemma D.2 and the relation between \(Y\) and \(W\) is from the definition of combine.

Next we show \(Y\) is \((3v)\)-swap distance from \(X\). By lemma D.2 there is a set of parameters \(\{(a_s, b_s, c_s, d_s, \epsilon_s)\}_{s \in [t]}\), where \(\sum_{s \in [t]} \epsilon_s = 3v\), such that \(A_s = \text{swap}(A_{s-1}, a_s, b_s, c_s, d_s, \epsilon_s)\) for \(s \in [t]\), and \(A_0 = Z, A_t = W\).

Let \(a'_s = a_s - \mathbb{I}(a_s > i)\) and \(c'_s = c_s - \mathbb{I}(c_s > i)\) for \(i \in [t]\), where \(\mathbb{I}(\cdot)\) is the indicator function. Let \(A'_0 = X\) and \(A'_s = \text{swap}(A'_{s-1}, a'_s, b'_s, c'_s, d_s, \epsilon_s)\) for \(s \in [t]\). Then we have \(A'_t = Y\), which implies \(Y\) is \((3v)\)-swap distance from \(X\).

We proved all the conditions of the lemma and we conclude the proof.

\[\square\]

### E Description and guarantees of roundiRow

Here we provide a description of the subroutine roundiRow.
The guarantees of the above algorithm are summarized in the lemma below.

**Lemma 4.11.** The algorithm roundiRow takes as inputs $Y \in \mathbb{R}^{\ell \times [0,k]}$, an column index $j \in [0,k]$ and a row index $i \in [\ell -1]$ such that: $Y^\top Y \in \mathbb{Z}_{\geq 0}^{[0,k]}$, $Y_{ij} \geq o$ and $\sum_{i'=1}^{i-1} Y_{i'j} + Y_{ij} - o \in \mathbb{Z}_{\geq 0}$, where $o = \lfloor Y^\top 1 \rfloor_o - \lfloor [Y^\top 1]_i \rfloor$. Outputs a matrix $X \in \mathbb{R}^{\ell \times [0,k]}$ such that,

- $X \leq Y$ and $\|X - Y\|_1 \leq 1$.
- $[X^\top]_{i'} = [Y^\top]_{i'}$ for all $i' \in [i-1]$, $[X^\top]_i \in \mathbb{Z}_{\geq 0}$, and $X^\top 1 \in \mathbb{Z}_{\geq 0}^{[0,k]}$.

**Proof.** We first show that $X \in \mathbb{R}^{\ell \times [0,k]}$ is a non-negative matrix. Note that $X_{i',j'} = Y_{i',j'} - o$ if $i' = i$, $j' = j$ or else $X_{i',j'}$ takes one of the values in the set $\{Y_{i',j'}, Y_{i',j'} - \min(Y_{i',j'}, \text{residue})\}$. As $Y_{ij} \geq o$, we get that $X_{i',j'} \geq 0$ for all $i' \in [\ell]$ and $j' \in [0,k]$. Furthermore, it is immediate that, $X \leq Y$.

Consider any $i' \in [1,i-1]$ and it is immediate that (Line 2) $[X^\top]_{i'} = [Y^\top]_{i'} \in \mathbb{Z}_{\geq 0}$. Furthermore, $[X^\top]_i = [Y^\top]_i - o = \lfloor [Y^\top 1]_i \rfloor \in \mathbb{Z}_{\geq 0}$.

Consider any $j' \in [0,k]$ such that $j' \neq j$ and note that $X_{i',j'} = Y_{i',j'}$ for all $i' \in [\ell]$. Therefore, we get that $[Y^\top 1]_{j'} = [Y^\top 1]_{j'} \in \mathbb{Z}_{\geq 0}$. Now consider the index $j$,

$$
\sum_{i'=1}^{\ell} X_{i',j} + 1 = \sum_{i'=1}^{i-1} X_{i',j} + \sum_{i'=i+1}^{\ell} X_{i',j} + 1 = \sum_{i'=1}^{i-1} Y_{i',j} + \sum_{i'=i}^{\ell} (Y_{i',j} - w_{i'}) + 1,
$$

$$
= \sum_{i'=1}^{\ell} Y_{i',j} - o - \sum_{i'=i+1}^{\ell} w_{i'} + 1
$$

The second and third equalities follow from Lines 6 and 8 of the algorithm respectively. In the following we show that, $\sum_{i'=i+1,\ell} w_{i'} = 1 - o$. To show this equality all we need is to show that residue $= 0$ at the end of the loop, which holds when $\sum_{i'=i+1,\ell} Y_{i',j} \geq 1 - o$. As $\sum_{i'=1}^{\ell} Y_{i',j} \in \mathbb{Z}_{\geq 0}$, $o \in (0,1)_{\mathbb{R}}$ and $z \coloneqq \sum_{i'=1}^{i-1} Y_{i',j} - o \in \mathbb{Z}_{\geq 0}$, we get that,

$$
\sum_{i'=i+1,\ell} Y_{i',j} - z = \sum_{i'=i}^{\ell} Y_{i',j} - z - o \text{ for some integer } z \in \mathbb{Z}_{\geq 0}.
$$

In the following we show that $\sum_{i'=i}^{\ell} Y_{i',j} - z \geq 1$. As $\sum_{i'=i}^{\ell} Y_{i',j} - z \in \mathbb{Z}_{\geq 0}$ all we need is to show that $\sum_{i'=i}^{\ell} Y_{i',j} - z > 0$. Suppose $\sum_{i'=i}^{\ell} Y_{i',j} - z = 0$, then $\sum_{i'=i+1,\ell} Y_{i',j} = \sum_{i'=i}^{\ell} Y_{i',j} - z - o = -o < 0$, which is a contradiction as $Y_{i',j'} \geq 0$ for all $i' \in [\ell], j' \in [0,k]$. Therefore we get that for all $j' \in [0,k]$,

$$
\sum_{i'=i}^{\ell} X_{i',j'} + 1 = \sum_{i'=i}^{\ell} Y_{i',j'}.
$$

(17)
As $X \in \mathbb{R}^{s \times [0,k]}_{\geq 0}$ and $\sum_{i' \in [t]} Y_{i'j} \in \mathbb{Z}_{\geq 0}$ (requirements of the lemma) we get that, $\sum_{i' \in [t]} X_{i'j} \in \mathbb{Z}_{\geq 0}$. Therefore,

$X^T \mathbf{1} \in \mathbb{Z}^{[0,k]}_{\geq 0}$.

In the remainder of the proof we show that $\|X - Y\|_1 \leq 1$. Recall, earlier we showed that: $X \leq Y$, $\|X^T \mathbf{1}\|_1 = \|Y^T \mathbf{1}\|_1$, for all $j' \neq j$ and $\|X^T \mathbf{1}\|_1 + 1 = \|Y^T \mathbf{1}\|_1$. Combining these inequalities together, we immediately get that, $\|X - Y\|_1 = 1$.

We proved all the conditions of the lemma and we conclude the proof.

\[\square\]

\section{Guarantees of swapmatrixround}

To prove Lemma 4.7 we show a stronger version of the lemma.

\textbf{Lemma F.1.} For any matrix $A \in \mathbb{R}^{s \times t}$ (s \leq t) that satisfies $A^T \mathbf{1} \in \mathbb{Z}^t_{\geq 0}$. In algorithm swapmatrixround, for all $r \in \{0\} \cup [t]$,

- $(D^{(r)} + A^{(r)})$ is $(3r)$-swap distance from $A$.
- $D_{ij}^{(r)} \geq 0, A_{ij}^{(r)} \geq 0$ for all $i \in [s]$ and $j \in [t]$, $[A^{(r)}]^T_i \in \mathbb{Z}^t_{\geq 0}$ for all $i \in [r]$, $(A^{(r)})^T \mathbf{1} \in \mathbb{Z}^{[0,k]}_{\geq 0}$ and $\|D^{(r)}\|_1 \leq r$.

\textbf{Proof.} We prove this lemma through induction.

\textbf{Base case i=0} In this case $A^{(0)} = A$, $D^{(0)} = 0$, we have that all the conditions of the lemma are immediately satisfied.

\textbf{Inductive step r} We assume that the lemma conditions hold for $r-1$ and prove for the case $r$. We use $Y^{(r)}$ to denote the value of $Y$ in algorithm swapmatrixround when the enumerator is $r$. We verify the properties claimed by the lemma one by one, and we use the properties related to partialRound and roundRow given by Lemma 4.10 and Lemma 4.11 respectively.

Since $(Y^{(r)}, j) = \text{partialRound}(A^{(r-1)}, r)$, $Y^{(r)}$ is 3-swap distance from $A^{(r-1)}$, which implies $(Y^{(r)} + D^{(r-1)})$ is 3-swap distance from $(A^{(r-1)} + D^{(r-1)})$. Note that $(A^{(r)} + D^{(r)}) = (Y^{(r)} + D^{(r-1)})$, and $(A^{(r-1)} + D^{(r-1)})$ is $(3(r - 1))$-swap distance from $A$, so we have $(D^{(r)} + A^{(r)})$ is $(3r)$-swap distance from $A$.

Since $A^{(r)} = \text{roundiRow}(Y^{(r)}, j, r)$, we have $A_{ij}^{(r)} \geq 0$ and $(Y^{(r)} - A^{(r)})_{ij} \geq 0$ for all $i \in [s]$ and $j \in [t]$. Furthermore, since $D_{ij}^{(r-1)} \geq 0$ for all $i \in [s]$ and $j \in [t]$, and $D^{(r)} = (Y^{(r)} - A^{(r)}) + D^{(r-1)}$, we have $D_{ij}^{(r)} \geq 0$ for all $i \in [s]$ and $j \in [t]$.

Since $(Y^{(r)}, j) = \text{partialRound}(A^{(r-1)}, r)$, we have $[Y^{(r)}]^T_i = [A^{(r-1)}]^T_i \in \mathbb{Z}^t_{\geq 0}$ for $i \in [r-1]$.

Since $A^{(r)} = \text{roundiRow}(Y^{(r)}, j, r)$, we have $[A^{(r)}]^T_i = [Y^{(r)}]^T_i \in \mathbb{Z}^t_{\geq 0}$ for $i \in [r-1]$, and $[A^{(r)}]^T_i \in \mathbb{Z}^t_{\geq 0}$ for $i = r$. As a result, $[A^{(r)}]^T_i \in \mathbb{Z}^t_{\geq 0}$ for $i \in [r]$.

Since $A^{(r)} = \text{roundRow}(Y^{(r)}, j, r)$, we have $(A^{(r)})^T \mathbf{1} \in \mathbb{Z}^{[0,k]}_{\geq 0}$.

Since $A^{(r)} = \text{roundRow}(Y^{(r)}, j, r)$, we have $\|Y^{(r)} - A^{(r)}\|_1 \leq 1$. Then we have $\|D^{(r)}\|_1 \leq \|Y^{(r)} - A^{(r)}\|_1 + \|D^{(r-1)}\|_1 \leq r$ because $D^{(r)} = (Y^{(r)} - A^{(r)}) + D^{(r-1)}$.

\[\square\]

\textbf{Corollary F.2.} For any matrix $A \in \mathbb{R}^{s \times t}$ (s \leq t) that satisfies $A^T \mathbf{1} \in \mathbb{Z}^t_{\geq 0}$. The algorithm swapmatrixround returns matrices $A'$ and $B$ such that,

- $A'$ is $O(s)$-swap distance from $A$. 

\section*{24}
We now construct a pseudo distribution \( \mathbf{A}' \), for all \( i \in [s] \) and \( j \in [t] \), \( \mathbf{B} \mathbf{1}^T \in \mathbb{Z}_{\geq 0}^s \), \( \mathbf{B}^T \mathbf{1}^T \in \mathbb{Z}_{\geq 0}^t \) and \( \| \mathbf{A}' - \mathbf{B} \|_1 \leq O(s) \).

\( \mathbf{A}' \) is \( O(s) \)-swap distance from \( \mathbf{A} \), and it implies \( \mathbf{A}' \mathbf{1}^T = \mathbf{A} \mathbf{1}^T \) and \( \mathbf{A}'^T \mathbf{1}^T = \mathbf{A}^T \mathbf{1}^T \).

We run \( O(s) \) times partialRound, roundiRow, and matrix addition and subtraction, so the running time is \( O(s^2 t) \).

We proved all the conditions of Lemma 4.7 and we conclude the proof.

**G Proof for Lemma 4.9**

**Proof.** Let \( \{ r_i \}_{i \in [\ell']} \) be the set of distinct probability values of distribution \( p \). Let \( X \in \mathbb{R}_{\geq 0}^{\ell' \times [k]} \) be the maximizer of the following optimization problem,

\[
\max_{Y \in \mathbb{R}_{\geq 0}^{\ell \times [k]}} g(Y).
\]

By Lemma 3.2, the maximizer \( X \) satisfies,

\[
C_\phi \cdot g(X) \geq \exp(-O(\min(k, \ell')) \log n) \mathbb{P}(p, \phi),
\]

In the following we consider multiple distributions with different probabilities to denote the probability values explicitly, we extend the notation of \( g(X) \) to \( g(X, r) \). Given the solution \( X \), we now maximize over the probability values, that is, solve,

\[
\max_{r' \in [0,1]^\ell} g(X, r') \equiv \sum_{i \in [\ell'], j \in [0,k]} (m_j \log r'^{ij} X_{ij} - X_{ij} \log X_{ij}) + \sum_{i \in [\ell']} [X1]_i \log [X1]_i,
\]

such that \( \sum_{i \in [\ell']} [X1]_i r'^{ij} = 1 \).

The optimum solution \( r^* \) satisfies \( \sum_{i \in [\ell']} [X1]_i r^* = 1 \) and,

\[
r^*_i = \frac{\sum_{j \in [0,k]} m_j X_{ij}}{[X1]^T_i \sum_{i' \in [\ell']} \sum_{j \in [0,k]} m_j X_{i'j}} = \frac{\sum_{j \in [0,k]} m_j X_{ij}}{n[X1]^T_i}.
\]

Furthermore,

\[
C_\phi g(X, r^*) \geq C_\phi g(X, r^') \geq \exp(-O(\min(k, \ell')) \log n) \mathbb{P}(p, \phi).
\]

Substituting values of \( r^* \) in \( g(X, r^*) \) we get,

\[
g(X, r^*) = n \sum_{i \in [\ell']} [X1]^T_i r^*_i \log r^*_i - \sum_{i \in [\ell'], j \in [0,k]} X_{ij} \log X_{ij} + \sum_{i \in [\ell']} [X1]_i \log [X1]_i
\]

We now construct a pseudo distribution \( q \) by rounding down the probability values into set \( R \) as follows: for each \( i \in [\ell] \), the number of elements with probability value \( r_i \) is equal to \( \sum_{i' \in [\ell']} \lfloor r_i \rfloor \cdot [X1]_i \), where \( \lfloor y \rfloor \) is defined as \( \max \{ x \in R \mid x \leq y \} \). Define a new solution \( X' \in \mathbb{R}^{\ell \times [k]} \) as follows: \( X'_{ij} \equiv \sum_{i' \in [\ell']} \lfloor r_i \rfloor \cdot [X1]_i \). Let \( \beta = \sum_{i' \in [\ell']} \lfloor r_i \rfloor \cdot [X1]_i \), and note that \( ||q||_1 = \beta \). Now note that the following inequality immediately holds,

\[
\mathbb{P}\left( q, \frac{\phi}{||q||_1} \right) = \mathbb{P}(q, \phi) \beta^{-n} \text{ and } \mathbb{P}(q, \phi) \geq C_\phi g(X', r).
\]

Now consider,

\[
\log g(X', r) = \sum_{i \in [\ell], j \in [0,k]} (m_j \log r_i X'_{ij} - X'_{ij} \log X'_{ij}) + \sum_{i \in [\ell']} [X1]_i \log [X1]_i,
\]

\[
\geq \sum_{i \in [\ell], j \in [0,k]} m_j \log \lfloor r_i \rfloor X_{ij} - \sum_{i \in [\ell'], j \in [0,k]} X_{ij} \log X_{ij} + \sum_{i \in [\ell']} [X1]_i \log [X1]_i.
\]
Consider the first term in the above expression,

$$\sum_{i \in [\ell'], j \in [0, k]} m_j \log [r_i^*]_{ij} R X_{ij} = \sum_{i \in [\ell'], j \in [0, k]} m_j \log \frac{|r_i^*|}{\beta} R X_{ij} + \log \beta \sum_{i \in [\ell'], j \in [0, k]} m_j X_{ij},$$

$$= \sum_{i \in [\ell'], j \in [0, k]} m_j \log \frac{|r_i^*|}{\beta} R X_{ij} + n \log \beta \quad (22)$$

To simplify the above expression we define \( \alpha_i \) as, \( r_i^* (1 + \alpha_i) = \frac{|r_i^*|}{\beta} \), then we get that,

$$\sum_{i \in [\ell']} r_i^* (1 + \alpha_i) [X^T]_i = \sum_{i \in [\ell']} \frac{|r_i^*|}{\beta} [X^T]_i = 1.$$\]

However, we also have that, \( \sum_{i \in [\ell']} \alpha_i r_i^* [X^T]_i = \sum_{i \in [\ell']} \alpha_i r_i^* = 0 \). Combining both we get that, \( \sum_{i \in [\ell']} \alpha_i r_i^* [X^T]_i = 0 \). Substituting this equality in Equation 22 we get,

$$\sum_{i \in [\ell'], j \in [0, k]} m_j \log [r_i^*]_{ij} R X_{ij} = \sum_{i \in [\ell'], j \in [0, k]} m_j \log \frac{|r_i^*|}{\beta} R X_{ij} + n \log \beta$$

$$= \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* (1 + \alpha_i) X_{ij} + n \log \beta$$

$$= \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* X_{ij} + \sum_{i \in [\ell'], j \in [0, k]} m_j \log (1 + \alpha_i) X_{ij} + n \log \beta$$

$$\geq \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* X_{ij} + \sum_{i \in [\ell'], j \in [0, k]} m_j (\alpha_i - \alpha_i^2) X_{ij} + n \log \beta$$

$$\geq \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* X_{ij} - \sum_{i \in [\ell'], j \in [0, k]} m_j \alpha_i^2 X_{ij} + n \log \beta \quad (23)$$

In the fourth inequality we used the inequality \( \log (1 + x) \geq x - x^2 \), when \( x \in (0, 1) \). In the last inequality we used \( \sum_{i \in [\ell']} \alpha_i r_i^* [X^T]_i = 0 \). Now note that each of the \( \alpha_i \) satisfy, \( |\alpha_i| \leq O(\alpha) \). Substituting it in the above expression we get,

$$\sum_{i \in [\ell'], j \in [0, k]} m_j \log [r_i^*]_{ij} R X_{ij} \geq \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* X_{ij} - \alpha^2 n + n \log \beta, \quad (24)$$

where in the above inequality we used \( \sum_{i \in [\ell'], j \in [0, k]} m_j X_{ij} = \sum_{j \in [0, k]} m_j \phi_j = n \). Combining the above inequality with Equation 21, we get,

$$\log g(X', r) \geq \sum_{i \in [\ell'], j \in [0, k]} m_j \log [r_i^*]_{ij} R X_{ij} - \sum_{i \in [\ell'], j \in [0, k]} X_{ij} \log X_{ij} + \sum_{i \in [\ell']} [X1]_i \log [X1]_i,$$

$$\geq \sum_{i \in [\ell'], j \in [0, k]} m_j \log r_i^* X_{ij} - \sum_{i \in [\ell'], j \in [0, k]} X_{ij} \log X_{ij} + \sum_{i \in [\ell']} [X1]_i \log [X1]_i - \alpha^2 n + n \log \beta,$$

$$= \log g(X, r^*) - \alpha^2 n + n \log \beta \quad (25)$$

Combining everything we get the following,

$$\log P(p, \phi) - \log P\left(\frac{q}{|q|_1}, \phi\right) = \log P(p, \phi) - \log P(q, \phi) + n \log \beta,$$

$$\leq O(\min(k, \ell') \log n) + \log(C_q g(X, r^*)) - \log(C_q g(X', r)) + n \log \beta,$$

$$\leq O(\min(k, \ell') \log n) + \alpha^2 n.$$
Proof for Lemma 4.6

We firstly show a special case of Lemma 4.6.

**Lemma H.1.** If $A' = \text{swap}(A, i_1, i_2, j_1, j_2)$, then $A' \in Z_R^{\phi,\text{frac}}$ and $g(A') \geq \exp(-4\epsilon \log n) g(A)$.

**Proof.**

Since $A' = \text{swap}(A, i_1, i_2, j_1, j_2)$, we have $A' \Rightarrow A$ and $A' \Rightarrow A'$. As $A \in Z_R^{\phi,\text{frac}}$, we immediately get that $A' \in Z_R^{\phi,\text{frac}}$ and we also have

$$g(A') = \exp \left( \frac{\sum_{i \in [t], j \in [0,k]} [C_{ij}A'_{ij} - A'_{ij} \log A'_{ij}]}{\exp \left( \sum_{i \in [t], j \in [0,k]} [C_{ij}A_{ij} - A_{ij} \log A_{ij}] + \sum_{i \in [t]} [A^1_i] \log [A^1_i] \right)} \right)$$

$$\geq \frac{1}{\exp \left( 2\epsilon \log(n+1) \right)} \geq \exp \left( -4\epsilon \log n \right),$$

where the last inequality holds because $\frac{\delta(x)}{\delta(x) + 1} = (x + 1)$ and

$$C_{i_1j_1} - C_{i_1j_2} - C_{i_2j_1} + C_{i_2j_2} = m_{j_1} r_{i_1} - m_{j_2} r_{i_1} - m_{j_2} r_{i_2} + m_{j_1} r_{i_2} \geq 0.$$

We conclude the proof.

**Proof for Lemma 4.6.**

Since $A'$ is x-swap distance from $A$, there exists a set of parameters denoted by $(i_1^{(s)}, i_2^{(s)}, j_1^{(s)}, j_2^{(s)}, \epsilon^{(s)})_{s \in [t]}$, where $\sum_{s \in [t]} \epsilon^{(s)} \leq x$, such that $A^{(s)} = \text{swap}(A^{(s-1)}, i_1^{(s)}, i_2^{(s)}, j_1^{(s)}, j_2^{(s)}, \epsilon^{(s)})$ for $s \in [t]$, where $A^{(0)} = A$, $A^{(t)} = A'$.

We apply Lemma H.1 on $A^{(i)}$ and $A^{(i-1)}$ for each $i \in [t]$, and we have

$$A^{(0)} \in Z_R^{\phi,\text{frac}} \Rightarrow A^{(1)} \in Z_R^{\phi,\text{frac}} \Rightarrow A^{(2)} \in Z_R^{\phi,\text{frac}} \Rightarrow \cdots \Rightarrow A^{(t)} \in Z_R^{\phi,\text{frac}},$$

$$g(A^{(t)}) \geq \exp \left( -4\epsilon^{(t)} \log n \right) g(A^{(t-1)}) \geq \cdots \geq \exp \left( -4 \left( \sum_{s \in [t]} \epsilon^{(s)} \right) \log n \right) g(A^{(0)}),$$

which implies $A' \in Z_R^{\phi,\text{frac}}$ and $g(A') \geq \exp \left( -O(x \log n) \right) g(A)$. We conclude the proof.