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Matroid Matching: The Power of Local Search

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Abstract

We consider the classical *matroid matching* problem. Unweighted matroid matching for linear matroids was solved by Lovász, and the problem is known to be intractable for general matroids. We present a PTAS for unweighted matroid matching for general matroids. In contrast, we show that natural LP relaxations that have been studied have an $\Omega(n)$ integrality gap and moreover, $\Omega(n)$ rounds of the Sherali-Adams hierarchy are necessary to bring the gap down to a constant.

More generally, for any fixed $k \geq 2$ and $\epsilon > 0$, we obtain a $(k/2 + \varepsilon)$-approximation for matroid matching in k-uniform hypergraphs, also known as the matroid k-parity problem. As a consequence, we obtain a $(k/2 + \varepsilon)$-approximation for the problem of finding the maximum-cardinality set in the intersection of k matroids. We also design a 3/2-approximation for the weighted version of a known special case of matroid matching, the *matchoid problem*.

1 Introduction

The matroid matching problem was proposed by Lawler as a common generalization of two important polynomial-time solvable problems: the non-bipartite matching problem, and the matroid-intersection problem (see [27]). Unfortunately, it turns out that matroid matching for general matroids is intractable and requires an exponential number of queries if the matroid is given by an oracle (see [30, 21]). This result can be easily transformed into a standard NP-completeness proof for a concrete class of matroids (see [38]). An important result of Lovász is that (unweighted) matroid matching can be solved in polynomial time for *linear matroids* (see [29]). There have been several attempts to generalize Lovász' result to the weighted case. Polynomial-time algorithms are known for some special cases (see [41]), but for general linear matroids there is only a pseudopolynomial-time randomized exact algorithm (see [7]).

In this paper, we revisit the matroid matching problem for general matroids. Our main result is that while LP-based approaches including the Sherali-Adams hierarchy fail to provide any meaningful approximation, a simple local-search algorithm gives a PTAS (in the unweighted case). This is the first PTAS for general matroid matching and to our knowledge also the first example of a problem where there is such a dramatic gap between the performance of the Sherali-Adams hierarchy and a simple combinatorial algorithm. We also provide approximation results for a generalization of the problem to hypergraphs; more details follow.

We assume familiarity with matroid algorithmics (see [38], for example) and approximation algorithms (see [43], for example). Briefly, for a matroid $\mathcal{M}$, we denote the ground set of $\mathcal{M}$ by $V = V(\mathcal{M})$, its set of independent sets by $\mathcal{I} = \mathcal{I}(\mathcal{M})$, and its rank function by $r_{\mathcal{M}}$. For a given matroid $\mathcal{M}$, the associated *matroid constraint* is $S \in \mathcal{I}(\mathcal{M})$ or equivalently $|S| = r_{\mathcal{M}}(S)$.

In the *matroid hypergraph matching problem*, we are given a matroid $\mathcal{M} = (V, \mathcal{I})$ and a hypergraph $G = (V, \mathcal{E})$ where $\mathcal{E} \subseteq 2^V$. Note that the vertex set of the hypergraph G and the ground set of the matroid $\mathcal{M}$ are the same. The goal is to choose a maximum-cardinality collection of disjoint hyperedges $E^* \subseteq \mathcal{E}$ in

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hypergraph G, such that the set of vertices covered by hyperedges in E^* is an independent set in matroid $\mathcal{M}$. If G is a graph, we obtain the classical matroid matching problem.

The matroid hypergraph matching problem generalizes several classical optimization problems, namely:

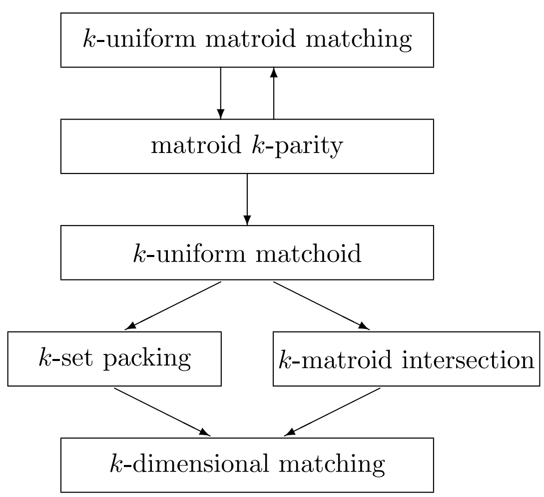
1. If M is a *free matroid* (i.e., $\mathcal{I}(\mathcal{M}) = 2^V$), then the problem is known as the *maximum hypergraph matching problem* or the *maximum set-packing problem.* Set packing in general is NP-hard, but when G is a graph (every hyperedge has exactly two vertices), it is the classical matching problem which led Edmonds to the very notion of polynomial-time algorithms (see [12, 13]).

2. In the *k-matroid intersection problem* we are given k matroids $\mathcal{M}_1 = (V, \mathcal{I}_1), \ldots, \mathcal{M}_k = (V, \mathcal{I}_k)$ on the same ground set V, and the goal is to find a maximum cardinality set S of elements that is independent in each of the k matroids, i.e. $S \in \cap_{j=1}^{k} \mathcal{I}_j$. The k-matroid intersection problem is NP-hard for $k \geq 3$ but polynomially solvable for $k = 2$ (see [38]).

3. A problem of intermediate generality is the *k-uniform matchoid* problem, defined for $k = 2$ by Edmonds and studied by Jenkyns (see [22]). In this problem, we have a k-uniform hypergraph and a matroid $\mathcal{M}_v$ given for each vertex v, having ground set the set of hyperedges containing v. The goal is to choose a maximum collection of hyperedges S, such that for each v, the hyperedges in S containing v form an independent set in $\mathcal{M}_v$. This can be also seen as a packing problem with many matroid constraints, where each item participates in at most k of them.

   By taking each $\mathcal{M}_v$ to be the uniform matroid of rank 1, we get the set-packing problem. By taking k arbitrary matroids defined on k copies of the same ground set V and a hypergraph of n parallel hyperedges on the k copies of the same element from V, we get k-matroid intersection. On the other hand, the matchoid problem is a special case of matroid matching, as we show below. We remark that even for $k = 2$, the matchoid problem is NP-hard (see [30]).

4. The special case of the matroid hypergraph matching problem when each vertex (i.e., element of the ground set) belongs to a unique hyperedge, and all hyperedges have cardinality exactly k is known as the *matroid k-parity problem*, or simply the *matroid parity problem* when $k = 2$. As we show below, this problem is in fact equivalent to k-uniform matroid matching, even in terms of approximation.



Next, we explain how the k-uniform matchoid problem is a special case of matroid k-parity. Given a hypergraph G, we can replace each vertex by n_v distinct copies, where n_v is the number of hyperedges containing v. We replace each hyperedge in G by a collection of distinct copies of its elements, so that we get a hypergraph G' where the hyperedges are disjoint. In the matchoid problem, we have a matroid $\mathcal{M}_v$ defined on the n_v copies of each vertex v, and we define a new matroid $\mathcal{M}'$ by taking the union of the matroids $\mathcal{M}_v$. Then matroid k-parity for $(G', \mathcal{M}')$ is equivalent to the original k-uniform matchoid for $(G, \mathcal{M}_v)$.

In fact, a similar construction implies that matroid k-parity includes (and therefore is equivalent to) matroid matching in k-uniform hypergraphs (or more generally in hypergraphs where each hyperedge has cardinality at most k, which can be shown by adding dummy elements). Given an instance of k-uniform matroid matching, we define n_v copies for each vertex v where n_v is the degree of v in the hypergraph $G = (V, \mathcal{E})$. We replace each hyperedge in G by a collection of distinct copies of its elements, so that the new hyperedges are disjoint. Let $G' = (V', \mathcal{E}')$ be the new hypergraph. We also define a new matroid $\mathcal{M}'$ on the ground set V', where the n_v copies of each vertex act as parallel copies. That is, a set of vertices $S' \subseteq V'$ is independent in $\mathcal{M}'$ if it contains at most one copy of each vertex from V and the respective set $S \subseteq V$ is independent in $\mathcal{M}$. It is not difficult to show that $\mathcal{M}'$ is a matroid.

Henceforth, we will discuss the matroid k-parity problem with the understanding that all our results can be easily extended to the k-uniform matroid matching problem. For the purposes of this paper, the terms *matroid matching* and *matroid parity* are essentially interchangeable.

Literature overview. There are a few different lines of research relevant to our results. The matroid parity problem (for $k = 2$) was originally popularized by Lawler [26]. The maximum-cardinality matroid parity problem was shown to have exponential query complexity in general (see [30, 21]), and to be NP-hard for some concrete classes of matroids. If the matroid is linear over the reals, the problem is polynomially solvable (see [29, 16, 34, 35, 40]). For more on matroid parity, applications and closely related problems, see [5, 31, 36, 37]. The linear matroid k-parity problem also admits a polynomial time algorithm if k is a constant, and the rank of matroid $\mathcal{M}$ is $O(\log |V|)$ (see [3]). The special case of matroid intersection can be solved in polynomial time for arbitrary matroids (see for example [38]), even in the weighted case.

It is not difficult to show that a simple greedy algorithm that adds hyperedges one-by-one, as long as the current solution remains independent, gives a k-approximation for general weighted matroid k-parity and that this guarantee is tight. This follows from the work of Jenkyns on k-independence systems; see [23] (see also Section 2). Until our present work, this was the only algorithm known for the general matroid k-parity problem with a provable approximation guarantee. The only improvement to our knowledge has been achieved in the case of unweighted matroid matching ($k = 2$), where Fujito proved that a local-search algorithm gives a 3/2-approximation (see [15]).

Linear-programming relaxations for the matroid matching problem ($k = 2$) have been studied in [42, 10, 11, 17]. An LP proposed by Vande Vate (see [42]) has been shown to be half-integral, and moreover, there are polynomial-time algorithms to find a half-integral optimal solution in the unweighted case (see [10, 11]) and the general weighted case (see [17]). However, this approach has not yielded any approximation algorithms for matroid matching.

We now survey known approximability and non-approximability results for special cases of the matroid k-parity problem. Local-search algorithms exploring a larger neighborhood for the k-set packing problem were analyzed in [20]. They showed that a local-search algorithm produces a solution with approximation guarantee $k/2+\varepsilon$ for any fixed $\varepsilon > 0$ and constant k in polynomial time. Hardness of approximation results are known for the maximum k-dimensional matching problem which is a special case of the maximum k-set packing problem and hence also matroid k-parity. The best known lower bound is the $\Omega(k/\log k)$ hardness of approximation of [19]. It is also known that a large-neighborhood local-search algorithm has a tight approximation guarantee of $k - 1 + \varepsilon$ for the problem of (weighted) k-matroid intersection (see [28]). It should be noted though that unweighted versions of packing problems seem to be easier for algorithm design and analysis, and approximation guarantees for general linear objective functions can be often improved for an unweighted variant of the problem.

The Sherali-Adams hierarchy (see [39]) has been studied recently for a number of combinatorial optimization problems (see [24, 25] for good surveys on the results in this area). Mathieu and Sinclair proved (see [32]) for the non-bipartite matching problem that r rounds of Sherali-Adams applied to the matching polytope have integrality gap $1 + O(1/r)$ and hence provides a PTAS (while the problem can be solved exactly in polynomial time). Chan and Lau considered k-uniform hypergraph matching, i.e. k-set packing, and proved that that even after $O(n)$ rounds of Sherali-Adams, the standard LP has integrality gap at least $k - 2$ (see [8]). In contrast, local search techniques yield a $(k/2 + \epsilon)$-approximation (see [20]), and an alternative LP (using intuition from local search) has integrality gap at most $(k + 1)/2$ (see [8]).

Our Results. On the negative side, we show that the known linear relaxations of matroid parity do not yield any reasonable approximation guarantee (even for $k = 2$ with unit weights). More precisely, all variants of the LPs that have been proposed have an $\Omega(n)$ integrality gap for instances with n pairs. Moreover, $\Omega(n)$ rounds of the Sherali-Adams hierarchy are required to generate an LP with a constant integrality gap.

In contrast, we prove that a very simple local-search algorithm gives a PTAS for unweighted matroid parity. Given the negative results for the matroid parity problem (see [30, 21]), this is the best type of worst-case result we could expect for this problem. It is also a strong manifestation of the fact that LP-based hierarchies do not always match the performance of combinatorial algorithms (which was, in a weaker sense, shown previously for the matching problem in graphs (see [32]) and hypergraphs (see [8])).

For the more general problem of unweighted matroid k-parity, we present a $(k/2 + \epsilon)$-approximation, for any fixed $k \geq 2$ and $\epsilon > 0$. As a special case, this subsumes the (unweighted) k-matroid intersection problem for which a k-approximation was known since 1976 (see [23]) and has been only recently improved to $k - 1 + \epsilon$ (see [28]).

The algorithm that we analyze is simple local search that in each iteration seeks to remove $s(\varepsilon)$ hyperedges and add $s(\varepsilon) + 1$ hyperedges to the current solution in such a way that the new solution defines an independent set in matroid $\mathcal{M}$. We call this the *s-neighborhood local-search algorithm*. If there is no improvement the algorithm stops and outputs the current local optimum. Our analysis uses an idea from [20] to reduce inductively the instance, and given the performance of the local search on the smaller instance, to derive the guarantee on the original one. But the presence of the matroid independence constraint complicates matters significantly. In particular, to achieve the approximation guarantee of $1 + \varepsilon$ for the matroid parity problem we need to implement the local-search algorithm with $s(\varepsilon)$ exponentially large in $1/\varepsilon$, while it is well known that $s(\varepsilon) = \lceil 1 + 1/\varepsilon \rceil$ is enough for the maximum matching problem in graphs (the fixed-size augmenting-path algorithm could be viewed as a local-search algorithm). We do not know at this point if having such a large neighborhood is necessary or if it is just an artifact of our analysis. Surprisingly, for $k \geq 3$, we show that to achieve $k/2 + \varepsilon$ approximation, it is enough to run the local-search algorithm with $s(\varepsilon)$ polynomially bounded in $1/\varepsilon$.

Finally, we give a 3/2-approximation algorithm for the weighted matchoid problem, which is a special case of weighted matroid parity. This result uses an LP relaxation of the matchoid problem and its known half-integrality (see [10, 11, 17, 42] for closely related linear programs). We provide an alternative proof that is very simple and intuitive which might be of independent interest.

The rest of the paper is organized as follows. In Section 2, we show that matroid k-parity is a special case of a "k-independence system" which implies a greedy k-approximation. In Section 3, we present our PTAS for unweighted matroid parity. In Section 4, we consider various linear-programming relaxations for the matroid parity problem and present our lower bounds on their integrality gap. In Appendix A, we present a $(k/2 + \varepsilon)$-approximation for matroid k-parity. In Appendix B, we present our 3/2-approximation for the weighted matchoid problem.

2 Relation to k-independence systems

First, we show that the matroid k-parity falls in the framework of k-independence systems (see [23]). Such systems generalize intersections of k matroids, and in fact several definitions of various degrees of generality have been proposed (see also [33, 6]). The definition of Jenkyns is as follows.

Definition 1 *For a family of sets $\mathcal{I} \subset 2^V$ and a set $W \subseteq V$, we define a* base *of W to be any inclusion-wise maximal subset $B \subseteq W$ such that $B \in \mathcal{I}$. We call $\mathcal{I}$ a p-*system*, if for any $W \subseteq V$,*

$$\max_{B:\text{base of } W} |B| \leq p \cdot \min_{B:\text{base of } W} |B|.$$

Lemma 1 *The independence system corresponding to matroid k-parity is a k-system.*

Proof: Consider an independent collection of hyperedges $W = \{e_1, \ldots, e_\ell\}$, and two bases B_1, B_2 of W. Assume toward a contradiction that $|B_2| > k|B_1|$. Let $S_1 = \bigcup\{e : e \in B_1\}$ and $S_2 = \bigcup\{e : e \in B_2\}$;

i.e. $|S_i| = k|B_i|$ and both S_1 and S_2 are independent in the matroid $\mathcal{M}$. By the matroid extension axiom, S_1 can be completed from S_2 to a set $S_1 \cup S'_2$ independent in $\mathcal{M}$, where $S'_2 \subseteq S_2 \setminus S_1$ and $|S'_2| = |S_2| - |S_1| = k|B_2| - k|B_1|$. Note that S'_2 is not necessarily a union of hyperedges. However, it must contain at least one hyperedge, otherwise $|S'_2| \leq (k-1)|B_2| < k|B_2| - k|B_1|$. Therefore, there is a hyperedge $e_i \in B_2 \setminus B_1$ that we can add to B_1 which contradicts B_1 being a base of W. □

The work of [22, 14] for p-systems gives the following results (see also [6]).

Theorem 1 *The greedy algorithm yields a p-approximation for maximizing any linear function over a p-system. Moreover, the greedy algorithm yields a $(p+1)$-approximation for the problem of maximizing a monotone submodular function over a p-system.*

Corollary 1 *The greedy algorithm yields a k-approximation for matroid k-parity, even in the weighted version. Moreover, the greedy algorithm yields a $(k+1)$-approximation for the problem of maximizing a monotone submodular function over sets feasible for the matroid k-parity problem.*

We regard the greedy k-approximation for matroid k-parity as a "folklore" result and a starting point for further improvements. For unweighted matroid parity ($k = 2$), this has been improved to a factor of 3/2 by Fujito [15]. For general k, no better approximation was known prior to our work.

3 PTAS for matroid parity

Let us start with the case of $k = 2$, i.e. matroid parity. In an instance of matroid parity, we have disjoint pairs, and we look for a maximum-cardinality collection of pairs whose union forms an independent set in a given matroid. We present a PTAS for this problem.

Definition 2 *For feasible solutions A and B of matroid parity, a "local move of size s between A and B" is a choice of $s-1$ pairs $e_1, \ldots, e_{s-1}$ inside A, and s pairs $e'_1, \ldots, e'_s$ inside B, such that $(A \setminus \bigcup_{i=1}^{s-1} e_i) \cup \bigcup_{i=1}^{s} e'_i$ is again feasible.*

Theorem 2 *For any $\epsilon > 0$, a local-search algorithm which considers local moves of size up to $s(\epsilon) = 5^{\lfloor 1/(2\varepsilon) \rfloor}$ achieves a $(1-\varepsilon)$-approximation for the matroid parity problem.*

The same result also holds for matroid matching, by a simple reduction that we outlined in the introduction. The theorem follows immediately from the following characterization of local optima.

Lemma 2 *Let $t \geq 1$, and A, B feasible solutions to the matroid parity problem such that*

$$|A| < \left(1 - \frac{1}{2t}\right)|B|.$$

Then there exists a local move of size 5^{t-1} between A and B.

Assuming that B is an actual optimum and A is a local optimum with respect to local moves of size 5^{t-1}, this implies that A is an $(1-1/2t)$-approximate solution. This means that for any fixed $\varepsilon > 0$, we can pick $t = \lceil 1/(2\varepsilon) \rceil$ and $s = 5^{t-1}$; the corresponding local-search algorithm achieves a $(1-\varepsilon)$-approximation for matroid parity.

It remains to prove the lemma. Our proofs uses the standard notion of matroid contraction. For a set $S \subset V(\mathcal{M})$, $\mathcal{M}/S$ (read $\mathcal{M}$ *contract* S) is the matroid having ground set $V(\mathcal{M}) \setminus S$ and set of independent sets $\{T \subseteq V(\mathcal{M}) \setminus S \ : \ T \cup J \in \mathcal{I}(\mathcal{M})\}$, where J is an arbitrary maximal independent subset of S with respect to $\mathcal{M}$.

Proof: Let A, B be feasible solutions as above. (We assume for simplicity that A and B are disjoint, otherwise we can contract the intersection, which only decreases the ratio $|A|/|B|$.) Because $|A| < |B|$, there exists $B_0 \subset B, |B_0| = |B| - |A|$ such that $A \cup B_0$ is independent in $\mathcal{M}$. We proceed by induction on t.

Base case: $t = 1$. For $t = 1$, we have $|A| < \frac{1}{2}|B|$. Then, $|B_0| = |B| - |A| > \frac{1}{2}|B|$. Because B decomposes into disjoint pairs, this means there must be a pair contained inside B_0. This pair can be added to A without violating independence, i.e. there is a local move of size one.

General case: $t \geq 2$. We assume that $|A| = |B| - a$ where $a > \frac{1}{2t}|B|$. We also assume $a \leq \frac{1}{2}|B|$, otherwise we are in the base case. We construct a set $B_0 \subset B$ as above, with $A \cup B_0$ independent and $|B_0| = a$. Again, if there is a pair contained inside B_0, we can add it to A, and we are done. So let us assume that no pair is contained completely inside B_0.

Every pair intersecting B_0 also contains an element in $B \setminus B_0$; let us denote the elements matched with B_0 by B_1. We have $|B_1| = |B_0| = a$. Let $\mathcal{M}_0 = \mathcal{M}/B_0$ denote the matroid where B_0 has been contracted. Because $A \cup B_0$ and $B_1 \cup B_0$ are independent in $\mathcal{M}$ (by construction), we get that A and B_1 are independent in $\mathcal{M}_0$. Because $|A| = |B| - a \geq a = |B_1|$, we can extend B_1 by adding (possibly zero) elements from A, to form an $\mathcal{M}_0$-independent set $(A \setminus A_1) \cup B_1$ where $|A_1| = |B_1| = a$.

If A_1 contains a pair e then we can find a local move as follows: $A \setminus e$ is independent in $\mathcal{M}_0$, and so is the set $(A \setminus A_1) \cup B_1$. Therefore, $A \setminus e$ can be extended to a set $(A \setminus e) \cup \{x', x''\}$ independent in $\mathcal{M}_0$, such that $x', x'' \in B_1$. The elements x', x'' are contained in pairs e', e'' whose remaining elements are in B_0. Because $(A \setminus e) \cup \{x', x''\}$ is independent in $\mathcal{M}_0 = \mathcal{M}/B_0$, any elements of B_0 can be added for free, and $(A \setminus e) \cup e' \cup e''$ is independent in $\mathcal{M}$. This defines a local move of size two.

The rest of the proof deals with the case when there is no pair contained in A_1. Then, every pair intersecting A_1 also contains an element in $A \setminus A_1$; let us denote the elements matched with A_1 by A_2. We have $|A_2| = |A_1| = a$. Here is where we apply the inductive hypothesis.

The inductive step. We define a new matroid $\mathcal{M}_1 = \mathcal{M}_0/B_1 = \mathcal{M}/(B_0 \cup B_1)$. By construction, the sets $A^* = A \setminus (A_1 \cup A_2)$ and $B^* = B \setminus (B_0 \cup B_1)$ are both independent in $\mathcal{M}_1$. They both form a union of pairs and hence are feasible solutions to the matroid parity problem for $\mathcal{M}_1$. We have $|A^*| = |A| - 2a$ and $|B^*| = |B| - 2a$. Because $|A| = |B| - a$, we get

$$\frac{|A^*|}{|B^*|} = \frac{|A| - 2a}{|B| - 2a} = \frac{|B| - 3a}{|B| - 2a} = 1 - \frac{1}{|B|/a - 2}.$$

Because we assumed $a > \frac{1}{2t}|B|$, we have $|B|/a < 2t$ and $|A^*| < (1 - \frac{1}{2t-2})|B^*|$, so we can apply the inductive hypothesis. There is a local move of size $s = 5^{t-2}$ between A^* and B^*, i.e. a union of $s - 1$ pairs $\tilde{A} \subseteq A^*$ and s pairs $\tilde{B} \subseteq B^*$ such that $(A^* \setminus \tilde{A}) \cup \tilde{B}$ is independent in $\mathcal{M}_1$. Our goal is to define a local move of size $5s = 5^{t-1}$ between A and B (in $\mathcal{M}$).

The set $(A^* \setminus \tilde{A}) \cup \tilde{B}$ is independent in $\mathcal{M}_1$. Unfortunately, $(A \setminus \tilde{A}) \cup \tilde{B}$ is not necessarily independent, even in $\mathcal{M}$. We have to proceed more carefully. The set $(A^* \setminus \tilde{A}) \cup A_2 = A \setminus (A_1 \cup \tilde{A})$ is independent in $\mathcal{M}_1 = \mathcal{M}_0/B_1$, because $(A \setminus A_1) \cup B_1$ was constructed to be independent in $\mathcal{M}_0$. Therefore, we can extend $(A^* \setminus \tilde{A}) \cup \tilde{B}$ to a set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ independent in $\mathcal{M}_1$, where $C_2 \subseteq A_2$ and $|C_2| \leq |\tilde{B}|$. (If $|A_2| \leq |\tilde{B}|$, we can just set $C_2 = A_2$.)

The new set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ is also independent in $\mathcal{M}_0$ (a weaker condition). So is $(A^* \setminus \tilde{A}) \cup (A_2 \setminus C_2) \cup A_1$, as any subset of A is independent in $\mathcal{M}_0$. Therefore, we can extend $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ to a set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2) \cup (A_1 \setminus C_1)$ in $\mathcal{M}_0$, where $C_1 \subseteq A_1$ and $|C_1| \leq |\tilde{B}|$.

The set we have obtained is not necessarily a union of pairs, so let us remove the whole pair for each element in C_1 and C_2. Let us denote by C' the union of all pairs intersecting $C_1 \cup C_2$. By our construction, we have $C_1 \cup C_2 \subseteq C' \subseteq A_1 \cup A_2$. Further, let us define $C'_1 = C' \cap A_1$ and $C'_2 = C' \cap A_2$. Each pair on $A_1 \cup A_2$ contains exactly one element in A_1 and one element in A_2, therefore $|C'_1| = |C'_2|$. Also, $|C'_1| = |C'_2| \leq |C_1 \cup C_2|$, because each each element of $C_1 \cup C_2$ contributes at most one pair to C'.

We obtain a feasible solution $A^+ = (A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C'_2) \cup (A_1 \setminus C'_1)$ in $\mathcal{M}_0$. Now, consider the set $(A^+ \setminus A_1) \cup B_1 = (A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C'_2) \cup B_1$. This is independent in $\mathcal{M}_0$, because $A^+ \setminus A_1$ was constructed to be independent in $\mathcal{M}_1 = \mathcal{M}_0/B_1$. Because A^+ misses some elements in A_1, namely C'_1, the cardinality of $(A^+ \setminus A_1) \cup B_1$ is actually larger than $|A^+|$, $|(A^+ \setminus A_1) \cup B_1| = |A^+| + |C'_1|$. Hence, we can extend A^+ by $|C'_1|$ elements of B_1, let us call them F_1, to obtain a set $A^+ \cup F_1$ independent in $\mathcal{M}_0$. The pairs touching F_1 have exactly 1 element in B_1 and the other element in B_0. Let F_0 be the elements of B_0 matched with F_1. We can add F_0 for free and obtain an independent set $A^+ \cup F_1 \cup F_0$ in $\mathcal{M}$. We have $|F_0| = |F_1| = |C'_1| = |C'_2|$. Now $A^+ \cup F_1 \cup F_0$ is a union of pairs and hence a feasible solution, of cardinality

$$|A^+ \cup F_1 \cup F_0| = |A^+ \cup C'_1 \cup C'_2| = |(A \setminus \tilde{A}) \cup \tilde{B}| > |A|.$$

Finally, let us estimate the size of this local move. We removed $\tilde{A} \cup C'_1 \cup C'_2$ from A, and added $\tilde{B} \cup F_1 \cup F_0$ instead. The size of C'_1 is bounded by $|C'_1| \le |C_1 \cup C_2| \le 2|\tilde{B}|$, hence $|F_1| = |C'_1| \le 2|\tilde{B}|$. The size of F_0 is equal to the size of F_1, i.e. $|F_0 \cup F_1| = 2|F_1| \le 4|\tilde{B}|$. In summary, we are adding at most $5|\tilde{B}|$ elements to A, i.e. the size of the local move is at most $5|\tilde{B}| = 5s = 5^{t-1}$. □

4 Linear-Programming Relaxations

In this section, we consider a linear-programming approach to matroid parity. Our results in this direction are mostly negative and indicate that linear programming in this case fails very badly compared to the local-search algorithm presented in the previous sections. We formulate our linear programs for the general case of matroid k-parity but the case $k = 2$ is already sufficiently general to obtain our results.

We start with the following natural LP for the weighted matroid k-parity problem (equivalent to an LP studied in [17]). The variables y_e correspond to the hyperedges and the variables x_u correspond to the elements of the ground set. We assume here that each hyperedge alone is independent; otherwise we remove it from the instance.

$$\max \sum_{e \in \mathcal{E}} w_e y_e, \tag{1}$$

$$\sum_{u \in S} x_u \le r_{\mathcal{M}}(S), \qquad \forall S \subseteq V, \tag{2}$$

$$x_u = y_e, \qquad \forall u \in e, e \in \mathcal{E}, \tag{3}$$

$$x_u, y_e \ge 0, \qquad \forall u \in V, e \in \mathcal{E}. \tag{4}$$

In the objective function (1) we are maximizing the total weight of chosen hyperedges. The constraints (3) correspond to the fact that we can choose a hyperedge only if we chose all its vertices. The constraints (2) are the standard rank constraints for the matroid $\mathcal{M}$ and any independent set of vertices must satisfy them.

For any set of elements $S \subseteq V$ let $\mathrm{sp}(S) = \{u \in V \mid r_{\mathcal{M}}(S \cup \{u\}) = r_{\mathcal{M}}(S)\}$ be the *span* of S in matroid $\mathcal{M}$. $F \subseteq V$ is a *flat* if $\mathrm{sp}(F) = F$ (for more information on these concepts, see [38], Chapter 39). Since $r_{\mathcal{M}}(\mathrm{sp}(S)) = r_{\mathcal{M}}(S)$, it is easy to see that it is enough to write the constraints (2) for every flat $F \subseteq V$; the inequality for arbitrary $S \subset V$ is implied by the flat $F = \mathrm{sp}(S)$.

Another set of valid inequalities (for $k = 2$) were suggested by Vande Vate [42] and studied in subsequent work [10, 11, 17, 42]. For a set $S \subseteq V$ and a hyperedge e, let $a(S, e) = r_{\mathcal{M}}(S \cap \mathrm{sp}(e))$. In case of a flat F, the intuition is that $a(F, e)$ is the dimension of the subspace of F generated by e. The LP proposed by Vande Vate is as follows.

$$\max \sum_{e \in \mathcal{E}} w_e y_e, \tag{5}$$

$$\sum_{e \in \mathcal{E}} a(S, e) y_e \le r_{\mathcal{M}}(S), \qquad \forall S \subseteq V, \tag{6}$$

$$y_e \ge 0, \qquad \forall e \in \mathcal{E}. \tag{7}$$

Again, it is equivalent to consider the inequalities (6) only for flats, which was the formulation given by Vande Vate. This LP is potentially stronger than LP (1-4), which can be equivalently obtained from (5-7) by replacing $a(S, e)$ with the smaller quantity $|S \cap e|$.

It is known that the linear program (5-7) is half-integral in the 2-uniform case. Moreover, there are polynomial time algorithms to find a half-integral optimal solution in the unweighted 2-uniform case (see [10, 11]) and the weighted 2-uniform case (see [17]). The following lemma shows the validity of the LP (5-7) in the general k-uniform case. Validity is not completely trivial, and we could not find a published proof of it (even for $k = 2$), so for completeness we provide a short proof.

Lemma 3 *The inequalities (6) are valid for the matroid k-parity problem.*

Proof: Consider any feasible solution, a collection of hyperedges $E^* = \{e_1, \ldots, e_k\}$ such that $e_1 \cup \ldots \cup e_k$ is an independent set in matroid $\mathcal{M}$. In the following, we denote the rank function of $\mathcal{M}$ simply by $r(S)$. Let $S_i = S \cap \mathrm{sp}(e_i)$. Note that $r(S_i) = a(S, e_i)$. We claim that for any $i < k$, $r(S_1 \cup \ldots \cup S_i) = r(S_1 \cup \ldots \cup S_{i-1}) + r(S_i)$. By induction, we will get that $r(S_1 \cup \ldots \cup S_k) = \sum_{i=1}^{k} r(S_i) = \sum_{i=1}^{k} a(S, e_i)$ which implies the Vande Vate constraint $\sum_{e \in E^*} a(S, e) \le r(S)$.

We let $r_A(S) = r(A \cup S) - r(A)$; due to submodularity, this is a non-increasing function of A. Let $A = S_1 \cup \ldots \cup S_{i-1}$ and $B = e_1 \cup \ldots \cup e_{i-1}$. Our goal is to prove that $r_A(S_i) = r(S_i)$. Since $A \subseteq \mathrm{sp}(B)$, we get $r_A(S_i) \ge r_{\mathrm{sp}(B)}(S_i) = r_B(S_i)$, using the fact that $r(\mathrm{sp}(B)) = r(B)$ and $r(S_i \cup \mathrm{sp}(B)) = r(S_i \cup B)$. On the other hand, as e_i is independent of B, we have

$$r(\mathrm{sp}(e_i)) = r_B(\mathrm{sp}(e_i)) = r_B(S_i) + r_{B \cup S_i}(\mathrm{sp}(e_i))) \le r_A(S_i) + r_{S_i}(\mathrm{sp}(e_i))$$

using again the submodularity of r. This implies that $r_A(S_i) \ge r(\mathrm{sp}(e_i)) - r_{S_i}(\mathrm{sp}(e_i)) = r(S_i)$. The opposite inequality is obvious and hence $r_A(S_i) = r(S_i)$. $\square$

In the following, we use examples where $e = \mathrm{sp}(e)$ for all hyperedges $e \in \mathcal{E}$. Note that in this case, $a(S, e) = r_{\mathcal{M}}(S \cap \mathrm{sp}(e)) = r_{\mathcal{M}}(S \cap e) = |S \cap e|$ and hence the two LPs are in fact equivalent.

4.1 Integrality gap example

It is known that the integrality gap of the linear-programming relaxation (1–4) is $k - 1 + \frac{1}{k}$ for the maximum weighted hypergraph matching problem [8]. Therefore, it is tempting to conjecture that a similar result should hold for matroid hypergraph matching. Unfortunately, as we show below, the integrality gap of the linear-programming relaxation (1–4) is $\Omega(|\mathcal{E}|)$ even when $k = 2$ and the matroid is linear over the rationals.

Example. Consider a ground set $V = \{u_1, v_1, \ldots, u_n, v_n\}$ of size $2n$, partitioned into pairs $e_i = \{u_i, v_i\}$. The weight of each pair is $w_{e_i} = 1$. Given an integer parameter $t \ge 1$, we define a matroid $\mathcal{M} = (V, \mathcal{I})$ as follows. For a set $S \subseteq V$, let $p(S)$ be the number of pairs e_i such that $e_i \subseteq S$. Then let $S \in \mathcal{I}$ if $p(S) \le t$.

It can be checked that $\mathcal{I}$ satisfies the matroid independence axioms: For any $S, T \in \mathcal{I}$, $|S| < |T|$, either T contains an element from a pair $\{u_i, v_i\}$ which is disjoint from S, or it contains more pairs than S. In either case, we can extend S by adding some element of T. Moreover, this matroid is is linear over the rationals (see Appendix C).

First, let us write down the LP for this particular example. We have variables y_i for $i = 1, \ldots, n$, which are constrained by $y_i \in [0, 1]$. Since $\mathrm{sp}(e_i) = e_i$, the LPs (1-4) and (5-7) coincide. It is enough to write the constraints (6) for flats, and in particular only for collections of pairs $S = \bigcup_{i \in T} e_i$. This is because including only one element of a pair in S always increases $r_{\mathcal{M}}(S)$ by 1 and hence cannot strengthen the constraint. Also, for $S = \bigcup_{i \in T} e_i$ where $|T| \le t$, the rank is $r_{\mathcal{M}}(S) = |S|$ and the respective constraint (6) is implied by $y_i \le 1$. The only non-trivial constraints are for $S = \bigcup_{i \in T} e_i, |T| > t$, where we get $r_{\mathcal{M}}(S) = 2t + (|T| - t) = t + |T|$. Also, $a(S, e_i) = 2$ for all $i \in T$. Therefore, the LP is as follows.

$$\max \sum_{i=1}^{n} w_i y_i, \tag{8}$$

$$\sum_{i \in T} y_i \le \frac{1}{2}(t + |T|), \qquad \forall T \subseteq [n], |T| > t \tag{9}$$

$$0 \le y_i \le 1, \qquad \forall i. \tag{10}$$

Lemma 4 *The integrality gap of LP (8-10) is $\Omega(n/t)$, even in the unweighted case.*

Proof: It is easy to see that $y_i = 1/2$ for all $i = 1, \ldots, n$ is a feasible fractional solution. Therefore, $LP \geq n/2$. However, only t pairs can be selected in an integral optimum, i.e. $OPT = t$. $\square$

For $t = 1$, we get an $\Omega(n)$ integrality gap. One way to improve the quality of linear-programming relaxations is to add valid inequalities that cut bad fractional solutions. One of the possible classes of valid inequalities are the so-called *clique inequalities* that were recently shown to reduce the integrality gap for unweighted hypergraph matching from $k-1$ to $(k+1)/2$ [8]. This motivates us to define the undirected graph $G' = (\mathcal{E}, E')$ where the vertices are the hyperedges $e \in \mathcal{E}$ in our instance of matroid hypergraph matching and the edges are defined between "incompatible hyperedges" e and e', i.e. when $r(e \cup e') < |e \cup e'|$. A set of vertices C in graph G' is called a clique if it has an edge between every pair of vertices in C. Let $\mathcal{C}$ be the set of all cliques in graph G'. Then the following set of constraints is valid for the matroid hypergraph matching problem

$$\sum_{e \in C} y_e \leq 1, \quad \forall C \in \mathcal{C}. \tag{11}$$

However, as we can see in the example above (for $t \geq 2$), sometimes the clique inequalities do not add any non-trivial constraints and the LP effectively remains the same. More generally, we could add all the valid constraints for the stable-set polytope corresponding to G' (or perhaps consider the semidefinite program corresponding to the Lovász θ-function). The relaxation would still remain the same, since the graph G' is empty in our example.

In the next section, we consider the strongest known systematic way of generating valid constraints in linear programming, which is the *Sherali-Adams hierarchy.*

4.2 The Sherali-Adams hierarchy

The Sherali-Adams hierarchy produces progressively stronger refinements of a given LP by introducing new variables y_L indexed by subsets of the original variables, and then projecting back to the space of the original variables. We follow the formalism of [32]. To carry out r rounds of Sherali-Adams, we consider all pairs of disjoint subsets of variables I, J such that $|I \cup J| = r$. We multiply each constraint by $\prod_{i \in I} y_i \prod_{j \in J}(1 - y_j)$, expand all the monomial terms and replace every square y_i^2 by y_i. Now all terms are multilinear, and we replace each occurrence of $\prod_{\ell \in L} y_\ell$ by a new variable y_L. We also do the same for the constraint $\prod_{i \in I} y_i \prod_{j \in J}(1 - y_j) \geq 0$, for all disjoint I, J such that $|I \cup J| = r + 1$. This defines the new LP; note that the variables y_L for $|L| > 1$ play no role in the objective function and thus the polytope can be viewed as projected back to the original space.

Lemma 5 *The integrality gap of LP (8-10) still remains $\Omega(n/r)$ after r rounds of the Sherali-Adams hierarchy.*

Proof: Our starting point is LP (8-10), with the parameter t chosen equal to the desired number of rounds r. We have constraints $\sum_{\ell \in T} y_\ell \leq \frac{1}{2}(r + |T|)$ for all $|T| > r$. We multiply this constraint by $\prod_{i \in I} y_i \prod_{j \in J}(1 - y_j)$ and obtain

$$\sum_{\ell \in T} y_\ell \prod_{i \in I} y_i \prod_{j \in J}(1 - y_j) \leq \frac{1}{2}(r + |T|) \prod_{i \in I} y_i \prod_{j \in J}(1 - y_j). \tag{12}$$

We expand the products, linearize the expressions, and replace monomials $\prod_{\ell \in L} y_\ell$ by new variables y_L as explained above. We also do the same thing for the constraints $\prod_{i \in I} y_i \prod_{j \in J}(1 - y_j) \geq 0$ with $|I \cup J| = r+1$. We claim that $y_L = 1/2^{|L|}$ for all $|L| \leq r + 1$ is a feasible solution for the new LP.

To see this, first observe that whenever we have $\ell \in J$ on the left-hand side of (12), the corresponding term contains $y_\ell(1 - y_\ell) = y_\ell - y_\ell^2$ which disappears after linearization. In terms where $\ell \in I$, we get y_ℓ^2 which gets linearized to y_ℓ. Equivalently, we can replace y_ℓ by 1 in its appearance before the product

$\prod_{i \in I} y_i$, whenever $\ell \in I$. Variables outside of $I \cup J$ remain unchanged. Therefore, after linearization, the left-hand side is equal to $(|T \cap I| + \sum_{\ell \in T \setminus (I \cup J)} y_\ell) \prod_{i \in I} y_i \prod_{j \in J}(1 - y_j)$.

Now we replace the monomials $\prod_{\ell \in L} y_\ell$ by y_L and substitute $y_L = 1/2^{|L|}$. Note that this is equivalent to directly substituting $y_\ell = 1/2$ for all ℓ. Thus the left-hand side becomes

$$(|T \cap I| + \frac{1}{2}|T \setminus (I \cup J)|)\, 2^{-|I \cup J|} = \frac{1}{2}(|T \cap I| + |T \setminus J|)\, 2^{-|I \cup J|} \leq \frac{1}{2}(r + |T|)\, 2^{-|I \cup J|}$$

using the fact that $|I| \leq r$. This verifies the linearized form of constraint (12).

The inequalities arising from $\prod_{i \in I} y_i \prod_{j \in J}(1 - y_j) \geq 0$ are easy to verify, since our assignment $y_L = 1/2^{|L|}$ is equivalent to substituting $y_i = 1/2$. Therefore, our fractional solution is feasible for r rounds of Sherali-Adams.

Finally, the value of our fractional solution is equal to $n/2$, because each singleton variable is $y_i = y_{\{i\}} = 1/2$. The integral optimum is $OPT = r$. □

To summarize, our LP (8-10) is an instance of the strongest "natural LP" for matroid matching we are aware of, namely the Vande Vate LP (5-7). The same LP (8-10) is obtained even with the added clique constraints (11) and other valid constraints for the stable-set polytope which are hard to optimize over in general. On top of this LP, we run the Sherali-Adams hierarchy and the gap still remains superconstant for $o(n)$ rounds.

5 Conclusion

We have seen that a simple combinatorial algorithm performs dramatically better than any known LP-based approach for matroid matching. Linear programming still holds some promise for the k-uniform matchoid problem. Theorem 4 and the results from [8] on the integrality gap of the hypergraph matching problem motivate the following.

Conjecture 1 *The integrality gap of the linear-programming relaxation (1–4) is $k-1+\frac{1}{k}$ for the maximum weighted k-matchoid problem and $k-1$ for the maximum weighted k-matroid intersection problem.*

In the case of weighted matroid k-parity, we have the following conjecture, which is true (and tight) for the weighted k-set packing problem due to [1] and also for the weighted k-matroid intersection problem due to [28].

Conjecture 2 *The simple local-search algorithm for the weighted matroid k-parity problem that tries to add/remove a constant number of hyperedges in each iteration has approximation guarantee $k - 1 + \varepsilon$ for any $\varepsilon > 0$ (with running time depending on $1/\varepsilon$).*

An especially intriguing open problem is to show that this simple local-search algorithm gives a PTAS for the weighted matroid parity problem ($k = 2$). This problem is interesting even for the special case of linear matroids, because Lovász' polynomial-time algorithm applies only to the unweighted case (see [29]). For the weighted linear case, there is only a pseudopolynomial-time randomized exact algorithm due to [7].

Another interesting line of research is to analyze more sophisticated local-search algorithms (see [4, 9]) implemented for the weighted matroid k-parity problem. Such algorithms are known to provide improved approximation guarantees for the weighted set packing problem.

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A Matroid k-parity

Here we extend the analysis of local search to matroid k-parity; i.e. instead of pairs, we work with hyperedges of size k. We assume here that $k \geq 3$. In an instance of matroid k-parity, all hyperedges are mutually disjoint. We remark that again, our analysis extends to k-uniform matroid matching where hyperedges are not necessarily disjoint, by a standard reduction.

Interestingly, the analysis for $k \geq 3$ is slightly different and the complexity of our $(2/k-\varepsilon)$-approximation for $k \geq 3$ has a much better dependence on ε than our PTAS for $k = 2$ (matroid matching). More precisely, while we need local moves of size exponential in $1/\varepsilon$ in order to achieve a $(1-\varepsilon)$-approximation for matroid matching, local moves of size polynomial in $1/\varepsilon$ are sufficient to achieve a $(2/k - \varepsilon)$-approximation for matroid k-parity. We do not know whether our analysis is optimal in terms of size of the local move.

Definition 3 *For feasible solutions A and B of matroid k-parity, a "local move of size s between A and B" is a choice of $s-1$ hyperedges $e_1, \dots, e_{s-1}$ inside A, and s hyperedges $e'_1, \dots, e'_s$ inside B, such that $(A \setminus \bigcup_{i=1}^{s-1} e_i) \cup \bigcup_{i=1}^{s} e'_i$ is again feasible.*

Theorem 3 *For any $k \geq 3$ and $\varepsilon > 0$, a local-search algorithm which considers local moves of size up to $s(\varepsilon) = \lceil 1/\varepsilon^3 \rceil$ achieves a $(2/k - \varepsilon)$-approximation for the matroid k-parity problem.*

This follows easily from the following characterization of local optima.

Lemma 6 *Let $k \geq 3$, $t \geq 1$, and A, B feasible solutions to the matroid k-parity problem such that*

$$|A| < \left(\frac{2}{k} - \frac{1}{(k-1)^t}\right)|B|.$$

Then there exists a local move of size $(2k+1)^{t-1}$ between A and B.

Note that in order to achieve a $(2/k - \varepsilon)$-approximation, it suffices to pick $t = \lceil \log_{k-1}(1/\varepsilon) \rceil$ and $s(\varepsilon) = (2k+1)^{t-1} \leq 1/\varepsilon^{\log_{k-1}(2k+1)}$. Then, if A is a local optimum and B is a global optimum, the lemma implies that $|A| \geq (2/k - 1/(k-1)^t)|B| \geq (2/k - \varepsilon)|B|$. For simplicity, we replaced $1/\varepsilon^{\log_{k-1}(2k+1)}$ by $1/\varepsilon^3$ in the statement of the theorem, but for large k the dependency gets close to $1/\varepsilon$.

It remains to prove the lemma.

Proof: Let A, B be feasible solutions as above. (We assume for simplicity that A and B are disjoint, otherwise we can contract the intersection, which only decreases the ratio $|A|/|B|$.) Because $|A| < |B|$, there exists $B' \subset B, |B'| = |B| - |A|$ such that $A \cup B'$ is independent in $\mathcal{M}$. We proceed by induction on t.

Base case: $t = 1$. Here, we have $|A| < (\frac{2}{k} - \frac{1}{k-1})|B| < \frac{1}{k}|B|$. Then, it is impossible that every hyperedge in B contains some element in $B \setminus B'$, because that would mean that $|A| = |B \setminus B'| \geq \frac{1}{k}|B|$. Hence, there must be a hyperedge contained completely inside B', which can be added to A without violating independence. This means there is a local move of size one.

General case: $t \geq 2$. We assume that $|A| = (2/k - \varepsilon)|B|$ and $\varepsilon > \frac{1}{(k-1)^t}$. Again, if there is a hyperedge contained inside B', we can add it to A, and we are done. So let us assume that no hyperedge is contained completely inside B'.

We use a counting argument to show that there must be many hyperedges with exactly $k-1$ elements in B'. Let a denote the number of such hyperedges ($|e \cap B'| = k-1$), and b the number of hyperedges such that $|e \cap B'| \leq k-2$. All hyperedges in B fall into one of these two categories, hence $|B| = k(a+b)$. On the other hand, $|B'| \leq (k-1)a + (k-2)b$ which means that $|A| = |B| - |B'| \geq a + 2b$. We assumed that $|A| = (2/k - \varepsilon)|B|$, which implies

$$a + 2b \leq |A| = \left(\frac{2}{k} - \varepsilon\right)|B| = (2 - k\varepsilon)(a+b). \tag{13}$$

We conclude that

$$a \geq k\varepsilon(a+b). \tag{14}$$

Let Q denote these a hyperedges in B and $V(Q)$ denote the elements of B that belongs to hyperedges in Q; each of them contains exactly one element in $B \setminus B'$ and $k-1$ elements in B'. Let $B_0 = V(Q) \cap B'$ and $B_1 = V(Q) \cap (B \setminus B')$. We have $|B_0| = (k-1)a$ and $|B_1| = a$.

Let $\mathcal{M}_0 = \mathcal{M}/B_0$ denote the matroid where B_0 has been contracted. Because $A \cup B_0 \subseteq A \cup B'$ and $B_1 \cup B_0 \subseteq B$, both of which are independent in $\mathcal{M}$, we get that A and B_1 are independent in $\mathcal{M}_0$. Because $|A| \geq a + 2b \geq |B_1|$, we can extend B_1 by adding (possibly zero) elements from A, to form a $\mathcal{M}_0$-independent set $(A \setminus A_1) \cup B_1$ where $|A_1| = |B_1| = a$.

If A contains any hyperedge e with $|e \cap A_1| \geq 2$, we find a local move of size two as follows: $((A \setminus e) \setminus A_1) \cup B_1$ is an independent set in $\mathcal{M}_0$, whose cardinality is at least $|A \setminus e| + 2$ (because A_1 contains ≥ 2 elements of e). Therefore, $A \setminus e$ can be extended to a set $(A \setminus e) \cup \{x', x''\}$ independent in $\mathcal{M}_0$, such that $x', x'' \in B_1$. The elements x', x'' are contained in hyperedges e', e'' whose remaining elements are in B_0. Because $(A \setminus e) \cup \{x', x''\}$ is independent in $\mathcal{M}_0 = \mathcal{M}/B_0$, any elements of B_0 can be added for free, and $(A \setminus e) \cup e' \cup e''$ is independent in $\mathcal{M}$. This defines a local move of size two.

The rest of the proof deals with the case that there is no hyperedge in A with more than 1 element in A_1. Let P be the collection of hyperedges in A intersecting A_1; each such hyperedge satisfies $|e \cap A_1| = 1$, and hence $|P| = |A_1| = a$. Let A_2 denote the remaining elements of P, i.e. $A_2 \subseteq A \setminus A_1$ and $|A_2| = (k-1)a$. Here is where we apply the inductive hypothesis.

The inductive step. We define a new matroid $\mathcal{M}_1 = \mathcal{M}_0/B_1 = \mathcal{M}/(B_0 \cup B_1)$. By construction, the sets $A^* = A \setminus (A_1 \cup A_2)$ and $B^* = B \setminus (B_0 \cup B_1)$ are both independent in $\mathcal{M}_1$. They both form a union of hyperedges and hence feasible solutions to the matroid k-parity problem for $\mathcal{M}_1$. We have $|A^*| = |A| - ka$ and $|B^*| = |B| - ka = kb$. Using (13), we get

$$\frac{|A^*|}{|B^*|} = \frac{|A| - ka}{kb} = \frac{(2 - k\varepsilon)(a+b) - ka}{kb} = \frac{2}{k} - \varepsilon - \frac{(k - 2 + k\varepsilon)a}{kb}$$

and applying (14) to estimate $a \geq kb\varepsilon$, we get

$$\frac{|A^*|}{|B^*|} \leq \frac{2}{k} - \varepsilon - (k - 2 + k\varepsilon)\varepsilon \leq \frac{2}{k} - (k-1)\varepsilon.$$

Because we assumed $\varepsilon > \frac{1}{(k-1)^t}$, we have $|A^*| < \left(\frac{2}{k} - \frac{1}{(k-1)^{t-1}}\right)|B^*|$, and we can apply the inductive hypothesis. There is a local move of size $s = (2k+1)^{t-2}$ between A^* and B^*, i.e. a union of $s-1$ hyperedges $\tilde{A} \subseteq A^*$ and s hyperedges $\tilde{B} \subseteq B^*$ such that $(A^* \setminus \tilde{A}) \cup \tilde{B}$ is independent in $\mathcal{M}_1$. Our goal is to define a local move of size $(2k+1)s$ between A and B (in $\mathcal{M}$).

We accomplish this by a construction essentially identical to the case of matroid parity. The set $(A^* \setminus \tilde{A}) \cup \tilde{B}$ is independent in $\mathcal{M}_1$. Unfortunately, $(A \setminus \tilde{A}) \cup \tilde{B}$ is not necessarily independent, even in $\mathcal{M}$. However, the set $(A^* \setminus \tilde{A}) \cup A_2 = A \setminus (A_1 \cup \tilde{A})$ is independent in $\mathcal{M}_1 = \mathcal{M}_0/B_1$, because $(A \setminus A_1) \cup B_1$ was constructed to be independent in $\mathcal{M}_0$. Therefore, we can extend $(A^* \setminus \tilde{A}) \cup \tilde{B}$ to a set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ independent in $\mathcal{M}_1$, where $C_2 \subseteq A_2$ and $|C_2| \leq |\tilde{B}|$. (If $|A_2| \leq |\tilde{B}|$, we just take $C_2 = A_2$.)

The new set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ is also independent in $\mathcal{M}_0$ (a weaker condition). So is $(A^* \setminus \tilde{A}) \cup (A_2 \setminus C_2) \cup A_1$, as any subset of A is independent in $\mathcal{M}_0$. Therefore, we can extend $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2)$ to a set $(A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2) \cup (A_1 \setminus C_1)$ independent in $\mathcal{M}_0$, where $C_1 \subseteq A_1$ and $|C_1| \leq |\tilde{B}|$.

The set we have obtained is not necessarily a union of hyperedges, so let us remove the entire hyperedge for each element in C_1 and C_2. Let us denote by C' the union of all hyperedges intersecting $C_1 \cup C_2$. Note that due to our construction, $C_1 \cup C_2 \subseteq C' \subseteq A_1 \cup A_2$. We also define $C'_1 = C' \cap A_1$ and $C'_2 = C' \cap A_2$. We know that each hyperedge on $A_1 \cup A_2$ contains exactly one element in A_1 and $k-1$ elements in A_2. Therefore, $|C'_2| = (k-1)|C'_1|$, and also $|C'_1| = \frac{1}{k}|C'| \leq |C_1 \cup C_2|$, because each element of $C_1 \cup C_2$ contributes at most one hyperedge to C'.

We obtain a feasible solution $A^+ = (A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2') \cup (A_1 \setminus C_1')$ in $\mathcal{M}_0$. Now, consider the set $(A^+ \setminus A_1) \cup B_1 = (A^* \setminus \tilde{A}) \cup \tilde{B} \cup (A_2 \setminus C_2') \cup B_1$. This is independent in $\mathcal{M}_0$, because $A^+ \setminus A_1$ was constructed to be independent in $\mathcal{M}_1 = \mathcal{M}_0/B_1$. Because A^+ misses some elements in A_1, namely C_1', the cardinality of $(A^+ \setminus A_1) \cup B_1$ is actually larger than $|A^+|$, namely $|(A^+ \setminus A_1) \cup B_1| = |A^+| + |C_1'|$. Hence, we can extend A^+ by $F_1 \subseteq B_1$, $|F_1| = |C_1'|$, to obtain a set $A^+ \cup F_1$ independent in $\mathcal{M}_0$. The hyperedges touching F_1 have exactly 1 element in B_1 and the remaining $k-1$ elements in B_0 (denote these by F_0), hence we can add F_0 for free and obtain an independent set $A^+ \cup F_1 \cup F_0$ in $\mathcal{M}$. We have $|F_1| = |C_1'|$ and $|F_0| = (k-1)|F_1| = (k-1)|C_1'| = |C_2'|$. To conclude, we have found a feasible solution $A^+ \cup F_1 \cup F_0$ in $\mathcal{M}$, of cardinality

$$|A^+ \cup F_1 \cup F_0| = |A^+ \cup C_1' \cup C_2'| = |(A \setminus \tilde{A}) \cup \tilde{B}| > |A|.$$

Finally, let us estimate the size of this local move. We removed $\tilde{A} \cup C_1' \cup C_2'$ from A, and added $\tilde{B} \cup F_0 \cup F_1$ instead. The size of C_1' is bounded by $|C_1'| \leq |C_1 \cup C_2| \leq 2|\tilde{B}|$, hence $|F_0 \cup F_1| = k|C_1'| \leq 2k|\tilde{B}|$. In summary, we are adding at most $(2k+1)|\tilde{B}|$ elements to A, i.e. the size of the local move is at most $(2k+1)|\tilde{B}| = (2k+1)^{t-1}$. □

B The weighted matchoid problem

In this section we analyze the integrality gap for the special case of the weighted matroid matching problem. First, we show a result similar to the one in [17].

Lemma 7 *The linear-programming relaxation (1–4) for the weighted matroid matching problem ($k = 2$) has a half-integral optimal solution, i.e. a solution x^*, y^* such that $x_u^*, y_e^* \in \{0, 1/2, 1\}$. Moreover, there exists a polynomial time algorithm to find such a solution.*

Proof: Consider the basic optimal solution (x^*, y^*) of the linear-programming relaxation (1–4). By the well-known properties of basic solutions, there exists a collection of sets $\mathcal{S}$ such that corresponding inequalities (2) are saturated, i.e. are satisfied with the equalities. Moreover, the solution (x^*, y^*) is a unique solution of the system of linear equations corresponding to these equalities and equations (3). Because the equations (3) are just equalities between different variables, there are $|\mathcal{E}| = |V|/2$ distinct variables and therefore we may assume that $|\mathcal{S}| = |\mathcal{E}|$.

By the well-known *uncrossing technique* we may assume that the sets in $\mathcal{S}$ form a laminar family, i.e. if $A, B \in \mathcal{S}$ then either $A \cap B = \emptyset$ or $A \subseteq B$ or $B \subseteq A$. Indeed, the submodularity of the rank function implies that the constraints corresponding to the sets $A \cap B$ and $A \cup B$ are also saturated. Therefore, if there are two sets $A, B \in \mathcal{S}$ such that $A \cap B \neq \emptyset$, $A \setminus B \neq \emptyset$ and $B \setminus A \neq \emptyset$ then we can replace A and B in collection $\mathcal{S}$ with $A \cup B$ and $A \cap B$.

Each laminar family can be viewed as a directed forest where the vertices correspond to sets in $\mathcal{S}$ and we connect two such vertices by a directed arc (A, B) if B is a minimal set in $\mathcal{S}$ such that $A \subseteq B$. Let E_F be the set of arcs in this directed forest and for each set $B \in \mathcal{S}$ let $\Upsilon(B) = \{A \in \mathcal{S} \mid (A, B) \in E_F\}$ be the set of children of B. The set of saturated constraints (2) corresponding to the set $\mathcal{S}$ can be rewritten as $\sum_{u \in B \setminus (\cup_{A \in \Upsilon(B)} A)} x_u = r_{\mathcal{M}}(B) - \sum_{A \in \Upsilon(B)} r_{\mathcal{M}}(A)$ for each $B \in \mathcal{S}$.

This new set of constraints has the crucial property that each variable x_u such that $x_u^* > 0$ appears in exactly one of the above constraints. Together with the constraints $x_u = x_v$ for $(u, v) \in \mathcal{E}$, we basically obtain an instance of the classical b-matching problem. We have a vertex corresponding to each $A \in \mathcal{S}$, $b(A) = r_{\mathcal{M}}(A) - \sum_{B \in \Upsilon(A)} r_{\mathcal{M}}(B)$ and edges correspond to pairs in the matroid matching instance (or constraints (3)). It is known that the b-matching polyhedron is half-integral [2], i.e. if we have variables defined on edges $y_e = x_u = x_v > 0$ and we know that this solution is a unique solution of the system of linear equations defined by the degree constraints in the b-matching problem then this solution must be half-integral. Therefore, we derive the statement of the lemma. □

Actually, the proof technique in the above lemma can be extended to the general hypergraph matroid matching problem. We can basically show that the properties of the basic solutions of the hypergraph matroid matching polyhedron are similar to the properties of the basic solutions of the hypergraph b-matching polyhedron. Because we do not know how to use such properties, we omit the statements and the proofs.

Let LP^* be the optimal value of the linear program (1–4). Using Lemma 7 we prove the following.

Theorem 4 *There exists a polynomial time algorithm that finds a feasible solution to the weighted matchoid problem of value at least $\frac{2}{3}LP^*$, given the half-integral optimal solution of the linear-programming relaxation (1–4).*

Proof: Recall the reduction of the matchoid problem to the matroid parity problem from Introduction. In this reduction each edge (u, v) in the matchoid problem corresponds to a pair that includes its own copies of u and v. Each vertex v in the matchoid instance has its own matroid $\mathcal{M}_v$ defined on edges incident to that vertex, or equivalently on copies of vertex v. The key property of the matchoid problem (or matroid parity instance corresponding to the matchoid problem) that allows us to derive an approximation algorithm is that all relevant rank constraints involve at most one element from each pair. This is because the rank constraints come from a matroid that is defined on the star of a vertex of the original instance.

Let (x^*, y^*) be the optimal half-integral solution of the linear-programming relaxation (1–4) defined on the instance of the matroid parity problem corresponding to the matchoid problem. Our algorithm rounds variables iteratively. In iteration i we are given an instance of the matroid parity problem corresponding to the matchoid problem. In this instance we have matroids $\mathcal{M}_{vi}$ for each vertex $v \in V$ of the original matchoid instance and we are given a set of pairs $\mathcal{E}_i$ such that the elements of the pair always participate in the different matroid constraints. We are also given a half-integral solution (x^i, y^i) of the linear program (1–4) for that instance. Initially, $\mathcal{E}_0 = \mathcal{E}$, $\mathcal{M}_{v0} = \mathcal{M}_v$ for each vertex $v \in V$ of the matchoid instance and $(x^0, y^0) = (x^*, y^*)$.

In iteration i our algorithm chooses a pair $e_h = (u, v) \in \mathcal{E}_i$ such that $y^i_{e_h} > 0$ of highest weight w_h. We add this pair to our current approximate solution and update matroids $\mathcal{M}_{vi+1} = \mathcal{M}_{vi}/e_h$ and $\mathcal{M}_{ui+1} = \mathcal{M}_{ui}/e_h$. For all other vertices $w \in V$, we define $\mathcal{M}_{wi+1} = \mathcal{M}_{wi}$. If $y^i_h = 1$ then we do not need to change the fractional solution. In this case $x^{i+1}_w = x^i_w$ and $y^{i+1}_{h'} = y^i_{h'}$ for all $w \neq u, v$ and $h' \neq h$. Otherwise, $y^i_h = 1/2$. In this case the old solution may not be feasible anymore for rank inequalities defined by matroids $\mathcal{M}_{ui+1}$ and $\mathcal{M}_{vi+1}$.

Note that x^i is a half-integral feasible solution of the matroid polyhedron $\mathcal{P}(\vee_{w \in V}\mathcal{M}_{wi})$ (rank constraints (2)) where $\vee_{w \in V}\mathcal{M}_{wi}$ the union of matroids $\mathcal{M}_{wi}$. By the integer decomposition property [38] (Corollary 42.1e, p.730) this vector is a convex combination of two integral independent sets I_1 and I_2 in this matroid (actually the proof will work almost the same way using any convex combination of independent sets). Each of those independent sets consists of a union of independent sets in matroids $\mathcal{M}_{wi}$, i.e. $I_1 = \cup_{w \in V} I_{w1}$ and $I_2 = \cup_{w \in V} I_{w2}$ where $I_{w1}, I_{w2} \in \mathcal{I}(\mathcal{M}_{wi})$.

Because $y_{hi} = x_{ui} = x_{vi} = 1/2$, we know that vertex u belongs to either I_{u1} or I_{u2}. Without loss of generality, we take $u \in I_{u1}$. By the matroid exchange properties there exists an element $\pi(u)$ such that the set $I_{u2} \cup \{u\} \setminus \{\pi(u)\} \in \mathcal{I}(\mathcal{M}_{ui})$, i.e. it is independent in this matroid. This element $\pi(u)$ belongs to some pair in $\mathcal{E}_i$ let us call it $e_{h'} = (\pi(u), w')$. If $y_{h'i} = x_{\pi(u)i} = x_{w'i} = 1/2$ then we define $y_{h'i+1} = x_{\pi(u)i+1} = x_{w'i+1} = 0$. If $y_{h'i} = x_{\pi(u)i} = x_{w'i} = 1$ then $y_{h'i+1} = x_{\pi(u)i+1} = x_{w'i+1} = 1/2$.

We apply an analogous operation for the element v, i.e. we define an element $\pi(v)$ and pair $e_{h''} = (\pi(v), w'')$ and update the variables accordingly. We do not change any other variables, i.e. $y_{hi+1} = y_{hi}$ for $h \neq h', h''$ and $x_{wi+1} = x_{wi}$ for $w \neq u, v, \pi(u), \pi(v), w', w''$.

We claim that the new fractional solution (x^{i+1}, y^{i+1}) is a feasible for the matroid matching problem for the next iteration. The key property we used is that the elements from the pair (u, v) participate in different matroid constraints and do not appear in the same independent set I_{ui} or I_{vi} for $i = 1, 2$.

In each iteration i we added a pair of value w_h to our approximate solution and decreased the value of the LP solution for the next iteration by at most $\frac{3}{2}w_h$. It implies that our final approximate solution will have value at least $\frac{2}{3}LP^*$. $\square$

We remark that the integrality gap of LP (1–4) is 3/2 even for the special case of non-bipartite matching (the example is a triangle), so with respect to this LP we cannot achieve a better approximation.

C Linearity of integrality gap example

Consider a ground set $V = \{u_1, v_1, \ldots, u_n, v_n\}$ of size $2n$, partitioned into pairs $e_i = \{u_i, v_i\}$. Given an integer parameter $t \geq 1$, we define a matroid $\mathcal{M}_t = (V, \mathcal{I})$ as follows. For a set $S \subseteq V$, let $p(S)$ be the number of pairs e_i such that $e_i \subseteq S$. Then let $S \in \mathcal{I}$ if $p(S) \leq t$.

For $t = 0$, the matroid $\mathcal{M}_0$ is a simple partition matroid — a set is independent if no more than one element is selected from each pair. For $t > 0$, the matroid $\mathcal{M}_t$ is related to $\mathcal{M}_0$ via a not very well known operation. Generally, for a matroid $\mathcal{M}$ of rank r on ground set V and an integer p satisfying $r \leq p \leq |V|$, the *elongation of* $\mathcal{M}$ *to height* p has as its family of bases the sets that have rank r in $\mathcal{M}$ and have cardinality p. Note that elongation is dual to the more familiar matroid operation of truncation (see [44, pp. 59–60]). The matroid $\mathcal{M}_t$ is the elongation of $\mathcal{M}_0$ to height $n + t$. Although there is a standard way of passing from a matrix representation of a matroid $\mathcal{M}$ to a matrix representation of an elongation of $\mathcal{M}$, the general procedure is not very parsimonious and it does not preserve the field that $\mathcal{M}$ is represented over (see [45, p. 402]). Instead, for $\mathcal{M}_t$, we give a direct parsimonious representation over the rationals.

Proposition 1 *The matroid $\mathcal{M}_t$ is linear over the rationals. Moreover the representation has bit size that is polynomial in n.*

Proof: Take W to be a $t \times n$ Vandermonde matrix. That is, we take distinct positive integers $c_1, c_2, \ldots, c_n$ (we can take $c_j = j$). Then, we fill column j of W with the t entries: $c_j^0, c_j^1, c_j^2, \ldots, c_j^{t-1}$. Our representation of the matroid $\mathcal{M}_t$ is

$$\left(\begin{array}{c|c} I & I \\ \hline W & 2W \end{array}\right)$$

where the first block of columns represents $u_1, \ldots, u_n$ and the second block represents $v_1, \ldots, v_n$.

Next, we consider any $S \subseteq V$ and we check that the respective columns of this purported representation are linearly independent if and only if S is independent in $\mathcal{M}_t$.

If S contains $t + 1$ pairs, we take the respective columns and delete all-zero rows. We obtain a $(t + 1 + t) \times 2(t + 1) = (2t + 1) \times (2t + 2)$ submatrix, which clearly has dependent columns as it has more columns than rows.

Suppose that S contains at most t pairs - we can assume it is a base which contains exactly t pairs and $n - t$ singletons. First, for each of the $n - t$ singletons, notice that the corresponding column has a non-zero in an associated row, that is zero in all of the other columns; so we can delete such columns and the associated rows, and concentrate on the columns corresponding to just the pairs. So we are led to a $2t \times 2t$ submatrix of the form

$$\left(\begin{array}{c|c} I & I \\ \hline A & 2A \end{array}\right),$$

where A is a $t \times t$ submatrix of W. We want to argue now that this $2t \times 2t$ matrix is non-singular. After elementary row operations, we pass to the matrix

$$\left(\begin{array}{c|c} I & I \\ \hline 0 & A \end{array}\right).$$

Now it is clear that we just need to check that A is non-singular. But the matrix A is itself a Vandermonde matrix, which is well known to be non-singular.

Finally, taking $c_j = j$, $j = 1, 2, \ldots, n$, the largest integer used is just $2n^{t-1}$, which has polynomial bit size. $\square$