Abstract. We consider the algorithmic decision problem that takes as input an \( n \)-vertex \( k \)-uniform hypergraph \( H \) with minimum codegree at least \( m - c \) and decides whether it has a matching of size \( m \). We show that this decision problem is fixed parameter tractable with respect to \( c \). Furthermore, our algorithm not only decides the problem, but actually either finds a matching of size \( m \) or a certificate that no such matching exists. In particular, when \( m = n/k \) and \( c = O(\log n) \), this gives a polynomial-time algorithm, that given any \( n \)-vertex \( k \)-uniform hypergraph \( H \) with minimum codegree at least \( n/k - c \), finds either a perfect matching in \( H \) or a certificate that no perfect matching exists.

1. Introduction

Matchings are fundamental objects in Graph Theory and have broad applications in other branches of Science and a variety of practical problems (e.g., the assignment of graduating medical students to their first hospital appointments\(^1\)). Applications of matchings in hypergraphs include the ‘Santa Claus’ allocation problem\(^3\); they also offer a universal framework for many important combinatorial problems, e.g., the Existence Conjecture for designs (see\(^{11, 22}\)) and Ryser’s conjecture\(^{34}\) on transversals in Latin squares.

This paper is concerned with the algorithmic question of finding a matching that is perfect, meaning that it covers all vertices of the graph or hypergraph. The graph case of this question is well understood: Tutte’s Theorem\(^{39}\) gives necessary and sufficient conditions for a graph to contain a perfect matching, and Edmonds’ Algorithm\(^{5}\) finds such a matching in polynomial time. However, for hypergraphs it is a different story: in fact, determining whether a 3-uniform hypergraph contains a perfect matching was one of Karp’s celebrated 21 NP-complete problems\(^{17}\). As the general problem is intractable (assuming \( \text{P} \neq \text{NP} \)), it is natural to seek conditions that guarantee a perfect matching, or at least make the existence question tractable.

1.1. Perfect matchings under minimum degree conditions. We start with some definitions that will be used throughout the paper. Given \( k \geq 2 \), a \( k \)-uniform hypergraph (in short, \( k \)-graph) \( H = (V, E) \) consists of a vertex set \( V \) and an edge set \( E \subseteq \binom{V}{k} \), where every edge is a \( k \)-element subset of \( V \). A matching in \( H \) is a collection of vertex-disjoint edges of \( H \). A perfect matching \( M \) in \( H \) is a matching that covers all vertices of \( H \). We always assume that \( k \) divides \( n := |V(H)| \), which is clearly a necessary condition for the existence of a perfect matching in \( H \). For \( S \subseteq V(H) \) the neighbourhood of \( S \) is \( N_H(S) := \{ T \subseteq V(H) \setminus S : S \cup T \in E(H) \} \), and the degree of \( S \) is \( \deg_H(S) = |N_H(S)| \); the subscript \( H \) is omitted if it is clear from the context. The minimum \( k \)-degree \( \delta_k(H) \) of \( H \) is the minimum of \( \deg_H(S) \) over all \( d \)-vertex sets \( S \) in \( H \). We refer to \( \delta_{k-1}(H) \) as the minimum codegree of \( H \).

Rödl, Ruciński and Szemerédi\(^{33}\) determined the sharp minimum codegree condition to ensure a perfect matching in an \( n \)-vertex \( k \)-graph for large \( n \) and all \( k \geq 3 \); the extremal examples are as follows.

\(^1\)In 2012, the Nobel Memorial Prize in Economics was awarded to Shapley and Roth “for the theory of stable allocations and the practice of market design.”
Construction 1.1. Given \( n \geq k \geq 3 \), let \( \mathcal{H}_{n,k} \) be the collection of \( k \)-graphs \( H \) such that there is a partition of \( V(H) = X \cup Y \) such that \( n/k - |X| \) is odd and all edges of \( H \) intersect \( X \) in an odd number of vertices.

Any matching \( M \) in \( H \in \mathcal{H}_{n,k} \) cannot be perfect; indeed, if it were we would have \( \sum_{e \in M} (|e \cap X| - 1) = |X| - n/k \), which is impossible, as the left hand side is even but the right hand side is odd. There is a large literature on minimum degree conditions that force a perfect matching, see [1, 4, 8, 12, 23–26, 29, 30, 32, 33, 36–38] and the surveys [31, 40], yet there are still many open problems, such as determining the minimum 1-degree condition that forces a perfect matching.

1.2. Algorithms. Let \( \text{DPM}_k(n, m) \) be the decision problem of determining whether an \( n \)-vertex \( k \)-graph \( H \) with \( \delta_{k-1}(H) \geq m \) contains a perfect matching. When can \( \text{DPM}_k(n, m) \) be decided in polynomial time? The result of [33] mentioned above shows that the decision problem is trivial for \( m \geq n/2 - k + 3 \) (then there is a perfect matching iff \( k \mid n \)). Szymańska [35] proved that if \( \delta < 1/k \) then \( \text{DPM}_k(n, \delta n) \) is polynomial-time reducible to \( \text{DPM}_k(n, 0) \), and so \( \text{NP} \)-complete. Karpinski, Rucinski and Szymańska [19] showed that there exists \( \varepsilon > 0 \) such that \( \text{DPM}_k(n, (1/2 - \varepsilon)n) \) is in \( \text{P} \) and posed the question of determining for the complexity of \( \text{DPM}_k(n, \delta n) \) for \( \delta \in (1/k, 1/2) \).

This was resolved for \( \delta > 1/k \) by Keevash, Knox and Mycroft [20, 21], who showed that the problem is in \( \text{P} \) and also gave a polynomial-time algorithm that either finds a perfect matching or a certificate that none exists. Their proof was very long and technical, and left open the ‘threshold case’ \( \delta = 1/k \), which poses additional challenges (discussed in Section 2). Han [14] completely resolved the question of [19] by showing that \( \text{DPM}_k(n, \delta n) \) is in \( \text{P} \) for \( \delta \in (1/k, 1/2) \). His proof was much simpler than that in [21] (it relied on some theory from [21] and also developed a lattice-based absorbing method which has found many other applications) but his result only concerned the decision problem, and left open the constructive problem, i.e. finding a perfect matching or a certificate that none exists.

Main results. The first contribution of this paper is a short proof of the following even stronger result.

Theorem 1.2. Let \( k \geq 3 \) and \( H \) be an \( n \)-vertex \( k \)-graph with \( \delta_{k-1}(H) \geq n/k - c \) for some \( c = o(n) \). Then there is an algorithm running in time \( n^{O(1)} 2^{O(c)} \) that finds a perfect matching in \( H \) or a certificate that none exists, where the implied constants depend on \( k \) only. In particular, if \( c = O(\log n) \) then \( \text{DPM}_k(n, n/k - c) \) is in \( \text{P} \).

In the conference version of this paper [16], we proved Theorem 1.2 with a worse running time, namely, \( O(n^{k \max\{4^k, c\}}) \), thereby obtaining a polynomial-time algorithm only when \( c \) is a constant. Our new improvement lies in the proof of the extremal case (Lemma 5.2), where the brute force search used in [16] is replaced by the colour-coding technique of Alon, Yuster and Zwick [2].

Our main result is the following algorithmic version of a theorem of Han [13] that \( \delta_{k-1}(H) \geq m \) implies a matching of size \( m \) when \( m < n/k \); our running time is fixed-parameter tractable with respect to \( c \), the deficiency in the minimum codegree condition.

Corollary 1.3. Let \( c, k, n \geq 3 \) and \( 0 < m \leq n/k \). Let \( H \) be an \( n \)-vertex \( k \)-graph with \( \delta_{k-1}(H) \geq m - c \). Then there is an algorithm running in time \( n^{O(1)} 2^{O(c)} \) that finds a matching of size \( m \) in \( H \) or a certificate that none exists, where the implied constants depend on \( k \) only.

The study of such algorithmic versions of classical problems in extremal graph theory was recently proposed by Fomin et al. [7], who obtained corresponding algorithmic versions of some extremal results for cycles in graphs, such as Dirac’s theorem. Our result can be viewed as an extension of a result of Fellows et al. [6] that covers the case \( m = c \). We note that Theorem 1.2 is the case \( m = n/k \). Furthermore, we will now show that the general statement is a simple consequence of this case.
Proof of Corollary 1.3. We start by considering the case that \( t := \frac{n-mk}{k-1} \in \mathbb{Z} \). Note that \( m + t = \frac{n + t}{k} \), so \( n + t \in k\mathbb{N} \). Consider any \( n \)-vertex \( k \)-graph \( H \) with \( \delta_k(H) \geq m - c \). We form a new \( k \)-graph \( H' \) with \( |V(H')| = n + t \) by adding a set \( A \) of \( t \) new vertices to \( V(H) \) and adding all \( k \)-vertex sets intersecting \( A \) to \( E(H) \). We note that \( \delta_{k-1}(H') \geq m - c + t \). Furthermore, writing \( \nu \) for the matching number, we have \( \nu(H') = \nu(H) + t \), as any matching in \( H \) can be greedily extended to a matching in \( H' \) with \( t \) extra edges, and from any matching in \( H' \) we can obtain a matching in \( H \) by deleting at most \( t \) edges that intersect \( A \). In particular, the existence of or finding matchings of size \( m \) in \( H \) is algorithmically equivalent to the same problem for perfect matchings in \( H' \). This can be solved by the algorithm given by Theorem 1.2, which finds a perfect matching in \( H' \) or a certificate that none exists. The running time is

\[
(n + t)^{O(1)}2^{O(c+k-1)} = n^{O(1)}2^{O(c)}.
\]

This proves the result if \( t := \frac{n-mk}{k-1} \in \mathbb{Z} \). For the general case, we choose \( s \leq k-1 \) such that \( n_0 := n - s \) has \( t := \frac{n_0-mk}{k-1} \in \mathbb{Z} \). We apply the above method to each of the \( \binom{n}{k} \) possible \( k \)-graphs \( H_0 \) on \( n_0 \) vertices obtained by deleting \( s \) vertices from \( H \), noting that \( \delta_{k-1}(H_0) \geq \delta_{k-1}(H) - s \geq m - c - (k - 1) \). \( \square \)

2. OVERVIEW OF THE PAPER

In the next section we describe the algorithm referred to in Theorem 1.2 and reduce its proof of correctness to two theorems (proved in the two subsequent sections) that respectively handle the ‘non-extremal’ and ‘extremal’ cases for \( H \). To explain this distinction, we consider the following construction that appears naturally around the codegree threshold \( n/k \).

Construction 2.1 (Space Barrier). Let \( V \) be a set of size \( n \) and fix \( S \subseteq V \) with \( |S| < n/k \). Let \( H \) be the \( k \)-graph on \( V \) whose edges are all \( k \)-sets that intersect \( S \).

We note that the minimum codegree of \( H \) is \( |S| \) and any matching in \( H \) has at most \( |S| < n/k \) edges, so cannot be perfect. On the other hand, Han [13] showed that any \( n \)-vertex \( k \)-graph with \( \delta_k(H) \geq n/k - 1 \) contains a matching of size \( n/k - 1 \), thus determining the tight codegree condition for a matching that is just one edge short of being perfect. This rather surprising phenomenon indicates that the key issue for whether there is a perfect matching near the codegree threshold \( n/k \) is whether \( H \) is close to a space barrier, or equivalently, whether \( H \) has an independent set of size about \( n - n/k \). We say \( H \) is \( \varepsilon \)-extremal if it has an independent set of size \( (1 - \varepsilon)\frac{n}{k}n \).

Our strategy of separating the non-extremal and extremal cases follows that of Han [14], and indeed several aspects of his proof carry over to our setting (despite our weaker assumption \( \delta_{k-1}(H) \geq n/k - c \), which is significantly more challenging to work with, as it does not rule out space barriers). However, the crucial difficulty that prevented Han from finding a perfect matching (as opposed to just testing for its existence) is the algorithmic intractability of finding (or even approximating) the largest independent set (this is NP-hard even in graphs, let alone hypergraphs).

The starting point for the new approach in this paper is to observe that this issue can be avoided via an algorithm of Han [13] that either finds an almost perfect matching or a large independent set. In the latter case, we can algorithmically extend to a large maximal independent set, which may not be of maximum size, but nevertheless gives us enough power to analyse the extremal case with some additional arguments. A final contribution of this paper (see the concluding remarks) is an algorithmic reduction of the perfect matching problem (see Problem 6.1) of independent interest.

3. THE ALGORITHM

In this section we state our algorithm and prove our main theorem assuming two theorems (concerning the non-extremal and extremal cases) whose proofs will be given in the following two sections.
The first ingredient of our algorithm is Procedure ListPartitions from [21, Section 2], which can test\(^2\) in time \(O(n^{k+1})\) whether \(H \in \mathcal{H}_{n,k}\) (defined in Construction 1.1). If \(H \in \mathcal{H}_{n,k}\) then the procedure finds a partition certifying that \(H\) does not have a perfect matching, so we reduce to the case \(H \notin \mathcal{H}_{n,k}\).

Next we require the following definitions from [21].

**Definition 3.1.** Let \(H = (V, E)\) be a \(k\)-graph and \(\mathcal{P} = \{V_1, V_2, \ldots, V_d\}\) be a partition of \(V\). The index vector of a set \(S \subseteq V\) is \(i_\mathcal{P}(S) = ([S \cap V_1], \ldots, [S \cap V_d]) \in \mathbb{Z}^d\). Given \(\mu > 0\), let \(I^\mu_\mathcal{P}(H)\) denote the set of all \(k\)-vectors \(i \in \mathbb{Z}^d\) such that at least \(\mu |V|^k\) edges \(e \in H\) have \(i_\mathcal{P}(e) = i\), and let \(L^\mu_\mathcal{P}(H)\) denote the lattice in \(\mathbb{Z}^d\) generated by \(I^\mu_\mathcal{P}(H)\).

The intuition behind \(L^\mu_\mathcal{P}(H)\) is that it captures the robust ‘divisibility’ constraints in \(H\) (generalising the parity condition in Construction 1.1) and that \(H\) should not have any divisibility obstruction to a perfect matching if it is possible to delete a small matching so that the index vector of the uncovered set is in \(L^\mu_\mathcal{P}(H)\).

We will also use the reachability methods introduced by Lo and Markström [27,28].

**Definition 3.2.** Let \(H\) be an \(n\)-vertex \(k\)-graph. We say that two vertices \(u\) and \(v\) are \((\beta, i)\)-reachable in \(H\) if there are at least \(\beta n^{ik-1}\) \((ik-1)\)-sets \(S\) such that both \(H[S \cup \{u\}]\) and \(H[S \cup \{v\}]\) have perfect matchings. We say that \(U \subseteq V(H)\) is \((\beta, i)\)-closed in \(H\) if any two vertices \(u, v \in U\) are \((\beta, i)\)-reachable in \(H\).

The following statement is a simplified form of [14, Lemma 2.5]. Throughout this paper, \(x \ll y\) means that for any \(y > 0\) there exists \(x_0 > 0\) such that for any \(x < x_0\) the subsequent statement holds.

**Lemma 3.3.** Let \(1/n \ll \beta, \mu \ll \gamma \ll 1/k\), where \(k \geq 3\) is an integer. Then for each \(k\)-graph \(H\) on \(n\) vertices with \(\delta_{k-1}(H) \geq n/k - \gamma n\), we can find in time \(O(n^{2k-1k+1})\) a partition \(\mathcal{P} = \{V_1, \ldots, V_d\}\) such that each \(V_i\) is \((\beta, 2^{k-1})\)-closed in \(H\) and has \(|V_i| \geq n/k - 2\gamma n\).

The following theorem handles the non-extremal case: given the partition \(\mathcal{P}\) from Lemma 3.3, it provides an algorithm that finds a perfect matching or outputs an independent set witnessing that \(H\) is in the extremal case.

**Theorem 3.4.** Suppose \(1/n \ll \beta, \mu \ll \gamma \ll 1/k\). Let \(H\) be a \(k\)-graph on \(n\) vertices such that \(\delta_{k-1}(H) \geq n/k - \gamma n\). Suppose \(\mathcal{P}\) is a partition of \(V(H)\) found by Lemma 3.3 such that there is a matching \(M_1\) with \(|M_1| \leq k\) and \(i_\mathcal{P}(V(H) \setminus V(M_1)) \in L_\mathcal{P}(H)\). Then there is an algorithm that finds in time \(O(n^{2k-1k+2})\) a perfect matching in \(H\) or an independent set in \(H\) of size \((1 - 5k\gamma)\frac{k-1}{k} n\).

The extremal case is handled by the following theorem (recall that \(H\) is \(\varepsilon\)-extremal if \(V(H)\) contains an independent subset of size at least \((1 - \varepsilon)\frac{k-1}{k} n\)).

**Theorem 3.5.** Assume \(1/n \ll \varepsilon \ll 1/k\) and \(c \leq \varepsilon n/k\). Let \(H\) be a \(k\)-graph on \(n\) vertices such that \(\delta_{k-1}(H) \geq n/k - c\). Suppose \(H \notin \mathcal{H}_{n,k}\) and \(H\) is \(\varepsilon\)-extremal. Then there is an algorithm that finds in time \(O(n^{4k-2O(\varepsilon n)})\) a perfect matching in \(H\) or a certificate that none exists.

Now we are ready to state our main algorithm (see Procedure PerfectMatching).

We conclude this section by showing correctness of this algorithm, thus proving Theorem 1.2, assuming Theorems 3.4 and 3.5, which will be proved in the following two sections. Suppose \(L\) is an edge-lattice in \(\mathbb{Z}^{|\mathcal{P}|}\), where \(\mathcal{P}\) is a partition of a set \(V\), then the coset group of \((\mathcal{P}, L)\) is \(G = G(\mathcal{P}, L) = L^{\mathcal{P}}_\text{max}/L\), where \(L^{\mathcal{P}}_\text{max} = \{x \in \mathbb{Z}^d : k \text{ divides } \sum_{i \in [d]} x_i\}\).

\(^2\) See [21, Lemma 2.2] and note that the proof is valid assuming \(\delta_{k-1}(H) \geq \omega n\) for any fixed \(\omega\) and \(n > n_0(k, \omega)\).
**Procedure** PerfectMatching

**Data** : an $n$-vertex $k$-graph $H$ such that $\delta_{k-1}(H) \geq n/k - c$.

**Output**: a perfect matching in $H$ or a certificate that none exists.

1. if Procedure ListPartitions finds that $H \in \mathcal{H}_{n,k}$ then
2. Output the certifying partition and halt.
3. Choose constants $1/n \ll \beta, \mu \ll \gamma \ll 1/k$;
4. Apply Lemma 3.3 to find a partition $P$ of $V(H)$ and $L_{P}^{\mu}(H)$;
5. if there is a matching $M_1$ of size at most $k$ with $i_{P}(V(H) \setminus V(M_1)) \in L_{P}^{\mu}(H)$ then
6. if Theorem 3.4 finds a perfect matching $M$ then
7. Output $M$ and halt.
8. else
9. Apply Theorem 3.5 with $\varepsilon = 5k\gamma$ and halt with appropriate output.
10. else
11. Output $\mathcal{P}$ as a certificate of “no perfect matching” and halt.

**Proof of Theorem 1.2.** We show correctness of Procedure PerfectMatching. As discussed above, Procedure ListPartitions from [21, Section 2], tests in time $O(n^{k+1})$ whether $H \in \mathcal{H}_{n,k}$ and if so outputs a partition certifying that $H$ does not have a perfect matching. Thus we can assume $H \notin \mathcal{H}_{n,k}$.

Next we apply Lemma 3.3 which finds $\mathcal{P} = (V_1, \ldots, V_d)$ in time $O(n^{2k+1}k^2)$; we note that $d \leq k$. Then we test each set of at most $k$ edges in $H$ (of which there are $O(n^2)$) to see if it is a matching $M_1$ satisfying $i_{P}(V(H) \setminus V(M_1)) \in L_{P}^{\mu}(H)$. If we find any such $M_1$ then we use it to apply Theorem 3.4, which finds a perfect matching or an independent set that can be used to apply Theorem 3.5, which in turn finds a perfect matching or a certificate that none exists. We only need to apply Theorems 3.4 and 3.5 for one such $M_1$ if it exists, so the running time is

$$O(\max\{n^{2k+1}k^2, O(n^{4k+O(kc)})\}) = O(n^{2k+1}k^2n^{O(kc)})$$

To complete the proof, it remains to show that if there is a perfect matching $M$ then some such $M_1$ exists. To see this we argue similarly to [21]. First we note the following property of $L_{P}^{\mu}(H)$ that follows easily from the minimum codegree condition of $H$: for any $v \in \mathbb{Z}^d$ with non-negative coordinates summing to $k - 1$ there is some $i \in [d]$ such that $v + u_i \in L_{P}^{\mu}(H)$. Indeed, consider all $(k - 1)$-sets with index vector $v$, by the minimum codegree condition $H$ contains at least $(n/k - 2\gamma n - k)k^{-1} \cdot \delta_{k-1}(H)/k! \geq (1/k - 3\gamma)k^k/k!$ edges whose index vector equals $v + u_i$ for $i \in [d]$; then by averaging, there exists $i \in [d]$ such that the number of edges with index vector $v + u_i$ is at least $\mu n k$. Then the proof of [21, Lemma 6.4] shows that the coset group $G$ of $L_{P}^{\mu}(H)$ in the lattice $\{v \in \mathbb{Z}^d : k | \sum v_i\}$ has size $|G| \leq d \leq k$. Now we repeatedly apply the pigeonhole principle to reduce $M$ to $M_1$ as in [21, Proposition 6.10]. We start with $M_1 = M$ and note that as $M$ is perfect we have $i_{P}(V(H) \setminus V(M)) = 0 \in L_{P}^{\mu}(H)$. While $|M_1| > k$ we consider any edges $e_1, \ldots, e_k$ in $M_1$ and the partial sums $\sum_{j=1}^{i} i_{P}(e_j)$ for $0 \leq i \leq k$. By the pigeonhole principle, some two of these sums lie in the same coset of $L_{P}^{\mu}(H)$, namely, there exist $0 \leq i_1 < i_2 \leq k$ such that

$$\sum_{j=i_1+1}^{i_2} i_{P}(e_j) = \sum_{j=1}^{i_2} i_{P}(e_j) - \sum_{j=1}^{i_1} i_{P}(e_j) \in L_{P}^{\mu}(H).$$

So we can delete $e_{i_1+1}, \ldots, e_{i_2}$ from $M_1$ while preserving $i_{P}(V(H) \setminus V(M_1)) \in L_{P}^{\mu}(H)$. We terminate with $|M_1| \leq k$, as required. \(\square\)
4. The non-extremal case

In this section we prove Theorem 3.4, which finds a perfect matching or a large independent set, thus establishing correctness of Procedure PerfectMatching in the non-extremal case. We adapt (and simplify) the approach of Han [14] via lattice-based absorption and also incorporate a derandomisation argument of Garbe and Mycroft [10] so that we can find a perfect matching (not just test for its existence).

4.1. Almost perfect matching or large independent set. The key idea of proof via absorption is that it simplifies the problem of finding a perfect matching to that of finding an almost perfect matching. Accordingly, we start by showing how to find an almost perfect matching or a large independent set. The following lemma is essentially [13, Lemma 1.6]; the proof is algorithmic, although this is not made explicit, so for the convenience of the reader we do this here.

Lemma 4.1. Suppose that $1/n \ll \gamma \ll 1/k$ and $k \mid n$. Let $H$ be a $k$-graph on $n$ vertices with $\delta_{k-1}(H) \geq n/k - \gamma n$. Then in time $O(n^{k+1})$ we can find either a matching that leaves at most $k^2/\gamma$ vertices uncovered, or an independent set of size $(1 - 2k\gamma)^{\frac{k-1}{k}} n$.

Proof. Consider any matching $M = \{e_1, e_2, \ldots, e_m\}$ in $H$. Let $V'$ be the set of vertices covered by $M$ and let $U$ be the set of vertices which are not covered by $M$. Assume that $|U| > k^2/\gamma$ and $U$ is an independent set (otherwise we can trivially enlarge $M$). We will show that we can find either a matching of size $m + 1$, or an independent set of size $(1 - 2k\gamma)^{\frac{k-1}{k}} n$.

We arbitrarily partition all but at most $k - 2$ vertices of $U$ into disjoint $(k - 1)$-sets $A_1, \ldots, A_t$ where $t = \lceil \frac{|U|}{k-1} \rceil > \frac{k}{\gamma}$. Let $D$ be the set of vertices $v \in V'$ such that $\{v\} \cup A_i \in E(H)$ for at least $k$ choices of $A_i$. First we consider the case that there is some $i^* \in [m]$ with $|e_{i^*} \cap D| \geq 2$. We fix distinct $x, y$ in $e_{i^*} \cap D$ and apply the definition of $D$ to pick distinct $A_i, A_j$ such that $\{x\} \cup A_i$ and $\{y\} \cup A_j$ are edges. Then we can enlarge $M$ by replacing $e_{i^*}$ by $\{x\} \cup A_i$ and $\{y\} \cup A_j$. Thus we may assume each $|e_i \cap D| \leq 1$.

Next we show that $|D| \geq \left(\frac{1}{k} - 2\gamma\right)n$. As $\delta_{k-1}(H) \geq n/k - \gamma n$ and $U$ is independent, 

$$t \left(\frac{1}{k} - \gamma\right)n \leq \sum_{i=1}^t \deg(A_i) \leq |D|t + n \cdot k < t(|D| + \gamma n),$$

so we have 

$$|D| > \left(\frac{1}{k} - 2\gamma\right)n.$$

Let $V_D := \bigcup\{e_i : e_i \cap D \neq \emptyset\}$. Note that $|V_D \setminus D| = (k - 1)|D| \geq (k - 1)\left(\frac{1}{k} - 2\gamma\right)n$. If $V_D \setminus D$ is independent then we can output it to conclude the proof. Thus we can assume there is some edge $e_0 \in H[V_D \setminus D]$. Suppose that $e_0$ intersects $e_{i_1}, \ldots, e_{i_l}$ in $M$ for some $l \leq k$ and write $\{v_{i_j}\} = e_{i_j} \cap D$ for all $j \in [l]$. By definition of $D$, we can greedily pick $A_{i_1}, \ldots, A_{i_l}$ such that $\{v_{i_j}\} \cup A_j \in E(H)$ for all $j \in [l]$. Then we can enlarge $M$ by replacing the edges $e_{i_1}, \ldots, e_{i_l}$ by $e_0$ and $\{v_{i_j}\} \cup A_j$ for $j \in [l]$.

For the running time, note that there are at most $n/k$ iterations in which we enlarge the matching. In each iteration we find $D$ by calculating the degree of each vertex $v$ in time $O(n^{k-1})$. Then we check if $V_D \setminus D$ is an independent set in time $O(n^k)$. The overall running time is $O(n^{k+1})$. \qed

4.2. An absorbing lemma. Now we implement the absorption part of the algorithm, Lemma 4.2, by derandomising a lemma from [14]. For the statement, we recall Definitions 3.1 and 3.2, and say that a set $T$ absorbs or is absorbing for a set $S$ if both $H[T]$ and $H[T \cup S]$ contain perfect matchings.
Lemma 4.2 (Absorbing Lemma). Let $1/n \ll 1/c' \ll \beta, \mu \ll 1/k, 1/t, 1/C'$. Suppose that $\mathcal{P} = \{V_1, \ldots, V_d\}$ is a partition of $V(H)$ such that each $V_i$ is $(\beta, t)$-closed in $H$. Then in time $O(n^{4tk^2})$ we can find a family $\mathcal{F}_{\text{abs}}$ of at most $c' \log n$ disjoint $tk^2$-sets such that $H[V(F)]$ contains a perfect matching for all $F \in \mathcal{F}_{\text{abs}}$ and every $k$-vertex set $S$ with $i_p(S) \in I_p^p(H)$ has at least $C'$ absorbing $tk^2$-sets in $\mathcal{F}_{\text{abs}}$.

The proof uses the following lemma which is [14, Claim 3.6].

Lemma 4.3. Suppose $V_i$ is $(\beta, t)$-closed in $H$ for all $i \in [d]$. Then any $k$-set $S$ with $i_p(S) \in I_p^p(H)$ has at least $\frac{\mu k^4}{2c'r^2} n^k$ absorbing $tk^2$-sets.

For the derandomisation we use the following lemma of Garbe and Mycroft [10, Proposition 4.7, Procedure SelectSet] (see also Karpiński, Ruciński and Szymańska [18]).

Lemma 4.4. Fix constants $\beta > \tau > 0$ and integers $m, M, N$ and $r \leq N$ such that $r$ and $N$ are sufficiently large, and that $M \leq (1/8) \exp(\tau^2 r/(3\beta))$. Let $U$ and $W$ be disjoint sets of sizes $|U| = M$ and $|W| = N$. Let $G$ be a graph with vertex set $U \cup W$ such that $G[U]$ is empty, $G[W]$ has precisely $\log r$ edges, and $\deg_G(u) \geq \beta N$ for every $u \in U$. Then in time $O(N^4 + Mn^3)$ we can find an independent set $R \subseteq W$ in $G$ such that $(1 - \nu)r \leq |R| \leq r$ and $|N_G(u) \cap R| \geq (\beta - \tau - \nu)r$ for all $u \in U$, where $\nu = 2mr/N^2$.

We are now ready to prove the absorbing lemma.

Proof of Lemma 4.2. We will apply Lemma 4.4 to the graph $G$ with parts $U = \{S : i_p(S) \in I_p^p(H)\}$ and $W = \{T \subseteq V(H) : |T| = tk^2\}$, where $T, T' \in W$ are adjacent iff $T \cap T' \neq \emptyset$, and $S \in U$ and $T \in W$ are adjacent iff $T$ absorbs $S$. In the notation of Lemma 4.4 we have $N = \binom{n}{tk^2}$, $M \leq \binom{n}{k}$ and $m = |E(G[W])| \leq tk^2 \binom{n}{tk^2-1} \binom{n}{tk^2} = t^2 k^4 N^2/(n - tk^2 + 1)$. We let $\beta' := \mu k^4/2k^2 + 1$, $\tau := \beta'/3$ and $r = c' \log n$. Then by Lemma 4.3, $\deg_G(u) \geq \beta' N$ for every $u \in U$, and

$$\exp\left(\frac{\tau^2 r}{3\beta'}\right) = \exp\left(\frac{\beta' c' \log n}{27}\right) \geq 8\binom{n}{k} \geq 8M,$$

as $n$ is large enough. Thus, in Lemma 4.4, in time $O(N^4 + MN^3) = O(n^{4tk^2})$, we can find a set $R \subseteq W$ which is independent in $G$ with $(1 - \tau)r \leq |R| \leq r$ and $|N_G(u) \cap R| \geq (\beta' - \tau - \nu)r$ for all $u \in U$, where

$$\nu = 2mr/N^2 \leq 2t^2 k^4 \log n/(n - tk^2 + 1)N^2 = 2t^2 k^4 c' \log n/(n - tk^2 + 1) < \beta'/3.$$

Note that $R$ consists of disjoint $tk^2$-sets of $H$ by definition of $G[W]$. We now remove $tk^2$-sets in $R$ that do not have a perfect matching, and denote the resulting family of $tk^2$-sets by $\mathcal{F}_{\text{abs}}$. Thus $|\mathcal{F}_{\text{abs}}| \leq c' \log n$, each member of $\mathcal{F}_{\text{abs}}$ has a perfect matching, and every $k$-vertex set $S$ with $i_p(S) \in I_p^p(H)$ has at least $\beta' r/3 \geq \beta' c' \log n/3 \geq C'$ absorbing $tk^2$-sets in $\mathcal{F}_{\text{abs}}$, as $n$ is large enough.

4.3. Proof of Theorem 3.4. We start the proof by fixing a parameter hierarchy $1/n \ll 1/c' \ll \beta, \mu \ll 1/k, 1/C \ll 1/k$, such that $C$ has the following property: for any $I \subseteq \{v \in \mathbb{Z}^d : \sum_i |v_i| \leq k\}$, where $d \leq k$, and $u$ in the integer span of $I$ with $\sum_i |u_i| \leq k^2(1 + \gamma^{-1})$, we can write

$$(4.1) \quad u = \sum_{v \in I} a_v(u)v.$$

where $a_v(u) \in \mathbb{Z}$ and $|a_v(u)| \leq C$ for each $v \in I$.

Let $H$ be a $k$-graph on $n$ vertices such that $\delta_{k-1}(H) \geq n/k - \gamma n$. Let $\mathcal{P} = \{V_1, \ldots, V_d\}$ be the partition found by Lemma 3.3; we note that $d \leq k$ by Lemma 3.3. Write

$$C' := (C + k + 2)^k + k/\gamma, \quad I = I_p^p(H), \quad \text{and} \quad L = L_p^p(H).$$
Let $M_1$ be a matching of size at most $k$ such that $\mathbf{i}_P(V(H) \setminus V(M_1)) \in L$. We first apply Lemma 4.2 to $H$ with $t = 2^{k-1}$, obtaining a family $\mathcal{F}_{abs}$ of $2^{k-1}k^2$ sets with $|\mathcal{F}_{abs}| \leq c' \log n$ such that every $S$ with $\mathbf{i}_P(S) \in I$ has at least $C'$ absorbing $2^{k-1}k^2$-sets in $\mathcal{F}_{abs}$. We obtain $\mathcal{F}_0$ from $\mathcal{F}_{abs}$ by removing any sets that intersect $V(M_1)$, of which there are at most $|V(M_1)| \leq k^2$; then every $S$ with $\mathbf{i}_P(S) \in I$ has at least $C' - k^2$ absorbing sets in $\mathcal{F}_0$. We let $M_0$ be a perfect matching on $V(\mathcal{F}_0)$ consisting of perfect matchings on each member of $\mathcal{F}_0$.

Next, we greedily find a matching $M_2$ in $V(H) \setminus V(M_0 \cup M_1)$ which contains $C$ edges $e$ with $\mathbf{i}_P(e) = i$ for every $i \in I$. This is possible because $H$ contains at least $\mu n^k$ edges for each $k$-vector $i \in I$, whereas $|V(M_0 \cup M_1 \cup M_2)| \leq 2^{k-1}k^2c' \log n + k^2 + kC(k^{d-1}) < \mu n$, as $n$ is large enough.

Let $H' := H[V(H) \setminus V(M_0 \cup M_1 \cup M_2)]$. Note that $|V(H')| \geq n - \mu n$ as $n$ is large, and

$$\delta_{k-1}(H') \geq \delta_{k-1}(H) - \mu n \geq n/k - 2\gamma n \geq (1/k - 2\gamma)|V(H')|.$$  

We apply Lemma 4.1 to $H'$ with parameter $2\gamma$ in place of $\gamma$. If Lemma 4.1 finds an independent set of size $(1 - 2k(2\gamma))\frac{k - 1}{k} |V(H')|$, then we can output this independent set and halt, as

$$(1 - 4k\gamma)\frac{k - 1}{k} |V(H')| \geq (1 - 4k\gamma)\frac{k - 1}{k} (n - \mu n) \geq (1 - 5k\gamma)\frac{k - 1}{k} n.$$  

Thus we can assume that Lemma 4.1 finds a matching $M_3$ covering all but a set $S_0$ of at most $k^2/(2\gamma) \leq k^2/\gamma$ vertices of $V(H')$.

Recall that $\mathbf{i}_P(V(H) \setminus V(M_1)) \in L$. By the definition of $M_2$, we have $\mathbf{i}_P(V(H) \setminus V(M_1 \cup M_2)) \in L$, which implies $\mathbf{i}_P(S_0) + \sum_{i \in H_0 \cup M_3} \mathbf{i}_P(e) \in L$. As in the previous section, by the pigeonhole principle we can find in time $O(n)$ edges $e_1, \ldots, e_{d'} \in M_0 \cup M_3$ for some $d' \leq r - 1$ such that $\mathbf{i}_P(S_0) + \sum_{i \in [d']} \mathbf{i}_P(e_i) \in L$.

We delete $e_1, \ldots, e_{d'}$ from our matching, thus leaving an unmatched set $D := \bigcup_{i \in [d']} e_i \cup S_0$. Then $\mathbf{i}_P(D)$ is in the integer span of $I$ and $|D| \leq k^2(1 + \gamma^{-1})$, so by definition of $C$, we have $\mathbf{i}_P(D) = \sum_{v \in I} a_v \mathbf{v}$ as in (4.1) with each $|a_v| \leq C$. We write each as $a_v = b_v - c_v$ such that one of $b_v, c_v$ is a nonnegative integer and the other is zero. Thus we have

$$(4.2) \quad \mathbf{i}_P(D) + \sum_{v \in I} c_v \mathbf{v} = \sum_{v \in I} b_v \mathbf{v}.$$  

Let $\mathcal{F} \subseteq M_2$ contain $c_v$ sets of index $v$ for each $v \in I$. We delete the family $\mathcal{F}$ of edges from $M_2$, thus leaving $V(\mathcal{F}) \cup D$ as the unmatched set.

The combinatorial meaning of (4.2) is that $V(\mathcal{F}) \cup D$ can be expressed as the disjoint union over $v \in I$ of $b_v$ sets of index $v$. Thus we regard $V(\mathcal{F}) \cup D$ as the union of $|\mathcal{F}| + |D|/k \leq kC + k + k/\gamma$ sets $S$ with $\mathbf{i}_P(S) \in I$. Each such $S$ has at least $C' - k^2 \geq d' + |\mathcal{F}| + |D|/k$ absorbing sets in $\mathcal{F}_0$, of which at most $d'$ may be unavailable due to deleting $e_1, \ldots, e_{d'}$, so we can greedily absorb all such $S$, thus obtaining a perfect matching of $H$.

For the running time, note that we can find $\mathcal{F}_{abs}$ in time $O(n^{2k+1}k^2)$, as $4tk^2 = 2^{k+1}k^2$. Then we find $M_2$ and $M_3$ both in time $O(n^{k+1})$, and $e_1, \ldots, e_{d'}$ in time $O(n)$. Finally, it takes constant time to find $b_v$ and $c_v$ for all $v \in I$, pick $\mathcal{F}$ and partition $V(\mathcal{F}) \cup D$; the absorption can also be done in constant time. The overall running time is $O(n^{2k+1}k^2)$.

5. The extremal case

In this section we prove Theorem 3.5, thus completing the proof of Theorem 1.2. We start by recalling the assumptions of the theorem, which will be assumed throughout the section. Let $k \geq 3$, suppose $\varepsilon > 0$ is sufficiently small, $n > n_0 = n_0(k, \varepsilon)$ is sufficiently large, and let $0 \leq c \leq \varepsilon n/k$ be an integer. Suppose $H$ is a $k$-graph on $n$ vertices with $\delta_{k-1}(H) \geq n/k - c$. Assume $H \notin \mathcal{H}_{n,k}$ and that $H$ is $\varepsilon$-extremal, i.e. there is an independent subset $S \subseteq V(H)$ with $|S| \geq (1 - \varepsilon)\frac{k - 1}{k} n$. We will
define an algorithm that finds either a perfect matching in \( H \) or a certificate that \( H \) has no perfect matching.

5.1. **Proof strategy and notation.** We start by introducing some notation and terminology, also used throughout the section, which we will motivate with reference to the strategy of the proof. We define a partition \((A, B, C)\) of \( V(H) \) as follows. Let \( C \) be any maximal independent subset of \( V(H) \) that contains \( S \). Note that one can construct \( C \) from \( S \) greedily in time \( O(n^k) \). For any \( x \in V \setminus C \), we write \( \text{deg}(x, C) \) for the number of edges \( e \) of \( H \) containing \( x \) with \( e \cap \{x\} \subseteq C \). We let

\[
A = \left\{ x \in V \setminus C : \text{deg}(x, C) \geq (1 - \alpha)\left(\frac{|C|}{k-1}\right) \right\},
\]

where \( \alpha = \varepsilon^{1/3} \). The final part of the partition is \( B = V(H) \setminus (A \cup C) \). We will see in Lemma 5.8 that \( B \) is small.

As \( C \) is independent, we have \( |e \cap C| \leq k - 1 \) for any edge \( e \), so we can assume \( |C| \leq (k - 1)n/k \), as otherwise it cannot be covered by a matching. Thus \( |C| \approx (k - 1)n/k \) and \( |A| \approx n/k \), so the bulk of any perfect matching can be thought of as a matching in an auxiliary graph between \( A' \subseteq A \) and a partition of \( C' \subseteq C \) into \((k - 1)\)-sets. The main task of the proof will be to efficiently check whether we can delete a small matching to reduce to such an auxiliary graph where a perfect matching can easily be found by derandomising a theorem of Pikhurko (see Theorem 5.9).

The **slack** of a matching \( M \) in \( H \) is defined as

\[
s_M := \left| (A \cup B) \setminus V(M) \right| - (n/k - |M|).
\]

As noted above, we can assume

\[
s := s_\emptyset = |A| + |B| - n/k = (k - 1)n/k - |C| \geq 0.
\]

In general, to use \( M \) as part of a perfect matching we require \( s_M \geq 0 \). Indeed, writing \( X = (A \cup B) \setminus V(M) \) and \( Y = C \setminus V(M) \), if \( s_M < 0 \) then \( |Y| = n - k|M| - |X| = (k - 1)|X| - ks_M > (k - 1)|X| \), so as \( Y \) is independent there cannot be a perfect matching.

We call a matching \( M \) a **cleaner** if \( M \) covers \( B \) and \( s_M \) is non-negative and even. Note that we impose the final condition to avoid a parity obstruction: we may only have edges that have odd intersection with \( X \), in which case a perfect matching of \( H[X \cup Y] \) requires \( |X| \) and \( n/k - |M| \) to have the same parity, namely, \( s_M \) is even. We summarise the above remarks as follows.

**Observation 5.1.** Suppose \( M \) is contained in a perfect matching and \( X = (A \cup B) \setminus V(M) \). Then:

(i) \( s_M \geq 0 \), and (ii) if all \( |e \cap X| \) with \( e \in E(H) \) are odd then \( s_M \) is even.

Next we identify a subset \( D \) of vertices in \( B \) with relatively high degree within \( C \); we will see in Lemma 5.11 below that it is easy to greedily extend any small matching to one covering \( D \). We list the vertices of \( B \) as \( v_1, \ldots, v_{|B|} \) so that \( \text{deg}(v_i, C) \) is a non-increasing sequence. We find (in linear time) the largest \( d > 0 \) such that \( \text{deg}(v_d, C) > (d + c)(k - 1)\left(\frac{|C|}{k-2}\right) \); if no such \( d \) exists then let \( d = 0 \). Let \( D := \{v_1, \ldots, v_d\} \) if \( d > 0 \), or \( D := \emptyset \) otherwise.

We conclude this subsection with some compact notation for describing the type of a set or edge with respect to the partition \((A, B, C)\) of \( V(H) \). We say that \( S \) is an \( A^iB^jC^l \) set if \( |S \cap A| = i \), \( |S \cap B| = j \) and \( |S \cap C| = l \); we also say \( S \) has the form \( A^iB^jC^l \). If any index is 0, we omit it, e.g. we write \( B^jC^l \) instead of \( A^0B^jC^l \). If \( S \) is an edge of \( H \) we call it an \( A^iB^jC^l \) edge. We also call it an \((i + j)\)-edge. We write \( N_H(v, A^iB^jC^l) \) for the \((k - 1)\)-sets in \( N_H(v) \) of form \( A^iB^jC^l \), namely,

\[
N_H(v, A^iB^jC^l) := \{S \in N_H(v) : |S \cap A| = i, |S \cap B| = j, |S \cap C| = l\}
\]

and also \( \text{deg}(v, A^iB^jC^l) = |N_H(v, A^iB^jC^l)| \).
5.2. The algorithm and proof modulo lemmas. Now we define our algorithm (it refers to some lemmas stated below). The notation is as in the previous subsection; we also let
\[ t = n/k - |A| = |B| - s \quad \text{and} \quad d = |D|. \]

**Procedure** PerfectMatchingEXT

**Data**: An \( \varepsilon \)-extremal \( n \)-vertex \( k \)-graph \( H \notin \mathcal{H}_{n,k} \) with \( \delta_{k-1}(H) \geq n/k - c \), an independent set \( S \) with \( |S| \geq (1 - \varepsilon)(k - 1)n \), and sets \( A, B, C, D \) as defined above.

**Output**: A perfect matching in \( H \) or a certificate that none exists.

1. if \( n < n_0 \) then
2. \( \quad \) Output “no PM” and \( C \), and halt.
3. if \( |C| > (k - 1)n/k \) then
4. \( \quad \) Output “no PM” and \( C \), and halt.
5. if \( s = |A| + |B| - n/k \) is odd and \( H \) contains no \( j \)-edge with \( j \leq s + 1 \) even then
6. \( \quad \) Output “no PM” and \( H \), and halt.
7. if there is no matching \( M_0 \) of size \( \max\{t - d, 0\} \) in \( H[(B \setminus D) \cup C] \) then
8. \( \quad \) Output “no PM” and \( (B \setminus D) \cup C \), and halt.
9. else
10. \( \quad \) Output “PM” ;
11. Use one of Lemmas 5.4, 5.5 or 5.6 to find a cleaner \( M \), then Lemma 5.3 to find a perfect matching.

Next we state five lemmas needed for the proof of Theorem 3.5 (the proofs are deferred to later in the section). The first lemma is required so that the search for \( M_0 \) is feasible, the second shows that it is sufficient to find a cleaner, and the others show how to find a cleaner in various cases.

**Lemma 5.2.** We have \( t - d < c \), and \( M_0 \) can be found in time \( O(n^k 2^{O(kc)}) \).

**Lemma 5.3.** If \( H \) has a cleaner \( M \) then \( H \setminus V(M) \) has a perfect matching, which can be found in time \( O(n^{4(k - 1)}) \).

**Lemma 5.4.** Suppose \( H \) contains no \( j \)-edge for all even \( 0 \leq j \leq k \) and \( H \notin \mathcal{H}_{n,k} \). If Procedure PerfectMatchingEXT outputs “PM” then we can find a cleaner in time \( O(n^k) \).

**Lemma 5.5.** Suppose \( s = 0 \) or that \( H \) has an \( i \)-edge for some even \( i \in [2, k] \) but does not have any \( ABC^{k-2} \) edge. If Procedure PerfectMatchingEXT outputs “PM” then we can find a cleaner in time \( O(n^k) \).

**Lemma 5.6.** Suppose \( s > 0 \) and \( H \) has some \( ABC^{k-2} \) edge \( e_0 \). If Procedure PerfectMatchingEXT outputs “PM” then we can find a cleaner in time \( O(n^k) \).

We conclude this subsection by assuming these lemmas and proving Theorem 3.5. First we make the following observation which will henceforth be used without comment.

**Observation 5.7.** If a matching \( M \) has \( n_i \) \( i \)-edges then \( s_M = s - \sum_{i \geq 1} n_i(i - 1) \).

**Proof of Theorem 3.5.** We need to show that Procedure PerfectMatchingEXT either finds a perfect matching in \( H \) or a certificate that none exists, in time \( \max\{O(n^4(k - 1)), O(n^{2k} 2^{O(kc)})\} = O(n^{4k} 2^{O(kc)}) \).

For the running time, we find \( C \) as an arbitrary maximal independent set containing \( S \) in time \( O(n^k) \), and then find \( A \) and \( B \) by determining their degrees to \( C \) in time \( O(n^k) \). Next we sort the
degrees of vertices from $B$ in time $O(n \log n)$ and find $d$ and $D$ in linear time. The first two tests runs in constant time and the third in time $O(n^6)$. By Lemma 5.2, the search for $M_0$ takes time $O(n^{2k_2}O(kc))$. If the output is "PM" then the conclusion of the algorithm finds a perfect matching in time $O(n^{4(k-1)})$ (assuming the lemmas).

It remains to show that there is no perfect matching if the algorithm outputs "no PM". The correctness of the first three tests follows from Observation 5.1. Indeed, for the third test, let $M = \{e\}$ for any $j$-edge $e$ with $j \geq s + 3$ even. Then $s_M = s - (j - 1) \leq -2$, so no matching containing $e$ is perfect. On the other hand, as $s$ is odd, any matching consisting of odd edges only cannot be perfect either. For the final test, note that in any matching $A \cup D$ can be incident to at most $|A \cup D| = n/k - t + d$ edges, so any perfect matching must contain a matching of size $\max\{t - d, 0\}$ completely within $(B \setminus D) \cup C$. 

5.3. Using a cleaner. In this subsection we prove Lemma 5.3, which shows how to find a perfect matching assuming that there is a cleaner. We start by estimating the sizes of the parts $A$, $B$, $C$; as discussed above, $B$ is small, $|A| = n/k$ and $|C| = n/n/k$.

**Lemma 5.8.** $|A| \geq n/k - \alpha^2 n$, $|B| \leq \alpha^2 n$ and $(1 - \varepsilon)(k-1)n/k \leq |C| \leq (k-1)n/k$.

*Proof.* The upper bound on $C$ follows from Observation 5.1 and the lower bound from our assumptions for Theorem 3.5, which give $|C| \geq (1 - \varepsilon)(k-1)n/k$. As $|A| + |B| + |C| = n$, we have

\[(5.1) \quad 0 \leq s = |A| + |B| - n/k = n/n/k - |C| \leq \varepsilon(k-1)n/k. \]

By the definitions of $A$ and $B$, we have

\[
\left(\frac{n}{k} - c\right)\left(\frac{|C|}{k-1}\right) \leq \sum_{x \in A \cup B} \deg(x, C) \leq (1 - \alpha)\left(\frac{|C|}{k-1}\right)|B| + \left(\frac{|C|}{k-1}\right)|A|. \]

We deduce $n/k - c \leq |A| + |B| - \alpha|B|$, so $\alpha|B| \leq |A| + |B| - n/k + c \leq \varepsilon n$ by (5.1) and $c \leq \varepsilon n/k$. Thus $|B| \leq \alpha^2 n$ and by (5.1) again, $|A| \geq n/k - |B| \geq n/k - \alpha^2 n$.

We also require the following algorithmic version of a special case of a result of Pikhurko [30], concerning perfect matchings in a $k$-graph $H$ that is $k$-partite, i.e. $V(H)$ has a partition $(V_1, \ldots, V_k)$ so that every edge intersects all $k$ parts. For $S \subseteq [k] = \{1, \ldots, k\}$ we write $\delta_S(H)$ for the minimum degree $\delta_H(\{v_i : i \in S\})$ of any set consisting of one vertex in each of the parts $(V_i : i \in S)$.

**Theorem 5.9.** [30, Theorem 3] Suppose $1/n \ll \gamma \ll 1/k$. Let $H$ be a $k$-partite $k$-graph with parts $V_1, \ldots, V_k$ each of size $n$. Suppose $\delta_{\{1\}}(H) \geq (1 - \gamma)n^{k-1}$ and $\delta_{[k] \setminus \{1\}}(H) \geq (1 - \gamma)n$. Then there is an algorithm that finds a perfect matching in $H$ in time $O(n^{4(k-1)})$.

As Pikhurko’s proof is probabilistic, we give an alternative derandomised proof via Lemma 4.4 (see also [15] for a similar proof).

*Proof.* We will apply Lemma 4.4 to the auxiliary graph $G$ with parts $U = V_1$ and $W = \prod_{i=2}^{k} V_i$, whose edges consist of all $\{u, S\}$ with $u \in U$, $S \in W$ and $\{u\} \cup S \in E(H)$, and all $\{S, S'\}$ with $S, S' \in W$ and $S \cap S' \neq \emptyset$. According to the notation of Lemma 4.4 we have $M = n$, $N = n^{k-1}$ and $m = |G[W]| \leq (k - 1)n^{2k-3}$. For any $u \in U$ we have $\deg(u, W) \geq \delta_{\{1\}}(H) \geq (1 - \gamma)n^{k-1}$, so we can take $\beta = 1 - \gamma$. To satisfy the conditions of Lemma 4.4 we let $\tau := 1/4$ and $r := 4\gamma n$. Then for large $n$ we have

\[
\exp\left(\frac{\tau^2 r}{3(1 - \gamma)}\right) \geq \exp\left(\frac{\gamma n}{12}\right) \geq 8n = 8M. \]
By Lemma 4.4, in time $O(N^4 + MN^3) = O(n^{4(k-1)})$, we can find an independent set $R$ in $G[W]$ such that $(1 - \nu)r \leq |R| \leq r$ and $|N_G(u) \cap R| \geq (\beta - \tau - \nu)r$ for all $u \in U$, where
\[
\nu = \frac{2mr}{N^2} \leq \frac{8(k-1)\gamma n^{2k-2}}{n^{2k-2}} = 8(k-1)\gamma.
\]
Thus $R$ is a collection of disjoint $(k-1)$-sets of the form $v_2 \ldots v_k$ with each $v_i \in V_i$ with $(1 - \nu)\gamma n \leq |R| \leq 4\gamma n$ and $|N_G(u) \cap R| \geq (1 - 8k\gamma - 1/4)r > r/2 = 2\gamma n$ for all $u \in V_1$, for small $\gamma$.

Next we partition $(\cup_{2 \leq i \leq k} V_i) \setminus V(R)$ arbitrarily into $n - |R|$ disjoint $(k-1)$-sets $S_1, S_2, \ldots, S_{n-|R|}$ each having one vertex in each of $V_2, \ldots, V_k$. We consider the bipartite subgraph $G'$ of $G$ induced by $V_1$ and $\{S_1, \ldots, S_{n-|R|}\}$.

We claim that the set $V'_1$ of vertices $v \in V_1$ with $\deg_{G'}(v) < (n - |R|)/2$ has size at most $2\gamma n$. To see this, first note that $|E(G')| \geq (1 - \gamma)n(n - |R|)$ as $\delta_{[k]\setminus\{1\}}(H) \geq (1 - \gamma)n$. Thus $G'$ has at most $\gamma n(n - |R|)$ non-edges, whereas every vertex in $V'_1$ is in at least $(n - |R|)/2$ non-edges of $G'$, so the claim holds.

We now proceed as follows.

1. Greedily match the vertices of $V'_1$ with $(k-1)$-sets in $R$. This is possible because each vertex $v$ forms an edge with more than $2\gamma n$ sets in $R$, and $|V'_1| \leq 2\gamma n$.
2. Greedily match the remaining sets of $R$ with vertices in $V_1 \setminus V'_1$. This is possible as $|R| \leq 4\gamma n$ and each set in $R$ forms an edge with at least $(1 - \gamma)n$ vertices in $V_1$.
3. Let $V''_1$ be the set of uncovered vertices in $V_1$ and $G''$ be the bipartite subgraph of $G$ induced by $V''_1$ and $\{S_1, \ldots, S_{n-|R|}\}$. We note that both parts of $G''$ have size $n - |R|$ and $G''$ has minimum degree at least $(n - |R|)/2$. Thus $G''$ has a perfect matching by Hall’s theorem, which can be found in time $O(n^3)$ by the Hungarian algorithm.

Clearly the union of the matchings above gives a perfect matching in $H$.

\[\square\]

**Proof of Lemma 5.3.** Suppose $H$ has a cleaner, i.e. a matching $M$ that covers $B$ such that $s_M = |(A \cup B) \setminus V(M)| - (n/k - |M|)$ is non-negative and even. We will show that $H \setminus V(M)$ has a perfect matching, which can be found in time $O(n^{4(k-1)})$.

We start by finding a matching $M_*$ of size $2\alpha^2 n$ with $V(M_*) \cap V(M) = \emptyset$ consisting only of $A^2C^{k-2}$ edges or only of $A^3C^{k-3}$ edges. To do so, we apply a greedy algorithm to choose edges one by one, which we call ‘good’ if they avoid $V(M)$ and all previous choices, until we obtain $2\alpha^2 n$ edges of one of the required forms. At each step, we consider any good set $S$ consisting of 2 vertices in $A$ and $k - 3$ vertices in $C$. As $\deg_H(S) \geq n/k - c > |B| + k|M| + 2k\alpha^2 n$ we can complete $S$ to a good edge of the form $A^2C^{k-2}$ or $A^3C^{k-3}$. Thus the algorithm to find $M_*$ can be completed, and clearly takes time $O(n^2)$.

We obtain $M'$ from $M$ by adding some edges of $M_*$: we add $s_M$ edges if they have the form $A^2C^{k-2}$ or $s_M/2$ edges if they have the form $A^3C^{k-3}$. In both cases we obtain $s_M' = 0$, i.e. $|C'| = (k-1)|A'|$ where $A' = A \setminus V(M')$ and $C' = C \setminus V(M')$. We partition $C'$ arbitrarily into $k-1$ parts $C^1, C^2, \ldots, C^{k-1}$ each of size $m := |A'|$. We note by (5.1) that $m \geq |A| - k|M| \geq |A| - k(|B| + s_M) > n/k - 2k\alpha^2 n$.

To find a perfect matching, it suffices to show that Theorem 5.9 applies to the $k$-partite sub-$k$-graph $H'$ of $H$ with parts $A', C^1, C^2, \ldots, C^{k-1}$.

To bound $\delta_{[k]\setminus\{1\}}(H')$ we consider any set $S$ formed by $k - 1$ vertices $v_i \in C^i$ for $i \in [k - 1]$. As $C$ is independent, the number of non-neighbours of $S$ in $A \cup B$ is at most
\[
|A| + |B| - n/k + c \leq \frac{(k-1)n}{k} + c \leq k\varepsilon m,
\]
where we use (5.1) in the first inequality and the last inequality follows from $m = |A'| \geq n/k - 2k\alpha^2 n > \frac{k-1}{k^2} n$. We deduce $\delta_{[k]\setminus\{1\}}(H') \geq m - k\varepsilon m = (1 - k\varepsilon)m$. 

\[\square\]
To bound \( \delta_{\{1\}}(H') \), we consider any \( v \in A' \subseteq A \), and note that by definition of \( A \), the number of \((k-1)\)-sets \( S \subseteq C \) with \( S \notin \mathcal{N}_H(v) \) is at most
\[
\alpha \frac{|C|^{k-1}}{(k-1)!} \leq \alpha \frac{(k-1)!}{(k-1)!} \leq \alpha \frac{(km)^{k-1}}{(k-1)!} = \alpha c_k m^{k-1},
\]
where \( c_k = \frac{k^{k-1}}{(k-1)!} \). This implies \( \delta_{\{1\}}(H') \geq (1 - \alpha c_k) m^{k-1} \).

By Theorem 5.9 with \( \gamma = \alpha c_k \), we find a perfect matching in \( H' \) in time \( O(n^{4(k-1)}) \).

5.4 Greedy extension. In this subsection we gather various results that will be used in the following subsection for greedy extension of matchings during the construction of a cleaner.

Lemma 5.10. Let \( B' \subseteq B \) and \( X \subseteq A \cup C \) with \(|X| \leq k|B| \). Then in time \( O(n^2) \) we can find a matching \( M_2 \) with \( V(M_2) \cap X = \emptyset \) that covers all vertices of \( B' \) by edges of the form \( ABC^{k-2} \) or \( BC^{k-1} \).

Proof. For each \( v \in B' \) in turn we pick \( k-2 \) arbitrary vertices from \( C \setminus X \), and an uncovered vertex in \( V \setminus (B \cup X) \) to complete an edge, which we add to \( M_2 \). Since \(|B| + |X| \leq 2k|B| < n/k-c \leq \delta_{k-1}(H) \), such an edge always exists.

Next we recall that \( D \) consists of \( d \) vertices each having degree in \( C \) larger than \((d+c)(k-1)\binom{|C|}{k-2} \). To see the following lemma, we consider a greedy algorithm for constructing \( D \), and note that at each step at most \((k-1)(d-1+c+1)\) vertices in \( C \) are already chosen, so the number of unavailable \((k-1)\)-sets in \( C \) is at most \((d+c)(k-1)\binom{|C|}{k-2} \).

Lemma 5.11. Given any set \( C_* \subseteq C \) of size at most \((c+1)(k-1) \), we can find a matching of \( M \) of size \( d \) such that \( D \in V(M) \subseteq D \cup (C \setminus C_*) \) in time \( O(n^k) \).

Now we give a degree bound that will be used to greedily cover \( B \setminus D \).

Lemma 5.12. If \( D \not\subseteq B \) then \( \deg(v, AC^{k-2}) > k|B||A|\binom{|C|}{k-3} \) for every vertex \( v \in B \setminus D \).

Proof. Suppose for contradiction that there is \( v \in B \setminus D \) with \( \deg(v, AC^{k-2}) \leq k|B||A|\binom{|C|}{k-3} \). By considering the degree of \( v \) together with each \((k-2)\)-set in \( C \) we have
\[
\deg(v, AC^{k-2}) + \deg(v, BC^{k-2}) + (k-1) \deg(v, C^{k-1}) \geq \binom{|C|}{k-2} \binom{n}{k} - c.
\]
As \(|B| \leq \alpha n^2 \) by Lemma 5.8 and \( \alpha \) is small we deduce
\[
(k-1) \deg(v, C^{k-1}) \geq \binom{|C|}{k-2} \binom{n}{k} - k|B|\binom{|C|}{k-3} - |B|\binom{|C|}{k-2} \geq k^2(|B| + c) - k|C| - k|C|.
\]
However, this contradicts the definition of \( D \), so the lemma holds.

We conclude this subsection by establishing the lower bound \( d = |D| \geq t-c \) and describing the algorithm to find \( M_0 \).

Proof of Lemma 5.2. Note that for any vertex \( b \in B \setminus D \), if \( d > 0 \) then \( \deg(b, C) \leq (d+c)(k-1)\binom{|C|}{k-2} \); if \( d = 0 \) then \( \deg(b, C) \leq (1+c)(k-1)\binom{|C|}{k-2} \). So it always holds that \( \deg(b, C) \leq (d+1+c)(k-1)\binom{|C|}{k-2} \). Recall that (i) \( \delta_{k-1}(H[B \cup C]) \geq t-c \) and independence of \( C \), and (ii) \( |B| \leq \alpha n^2 \) and \(|C| \leq (k-1)n/k \), we have
\[
\binom{|C|}{k-1}(t-c) \leq \sum_{b \in B} \deg(b, C) \leq d \binom{|C|}{k-1} + |B|(d+1+c)(k-1)\binom{|C|}{k-2} \leq (d+\alpha)\binom{|C|}{k-1}.
\]
This implies that \( d \geq t-c \) as \( t, c, d \in \mathbb{N} \).
PerfectMatchingEXT

Now assume $t - d > 0$, and we need to find a matching $M_0$ of size $t - d \leq c$ in $H[(B \setminus D) \cup C]$, if one exists. A greedy algorithm finds such a matching in time $O(n^{ck})$. Next we show how to improve the running time by the colour-coding technique by Alon, Yuster and Zwick [2]. This requires an explicit construction of a family of perfect hash functions, such as that due to Fredman, Komlós and Szemerédi [9]. For any $n, k' \in \mathbb{N}$ and set $U$ of size $n$, this gives a family $\mathcal{F}$ of functions $f : U \to X$ with $|X| = O(k')$ such that for any $k'$-subset $U'$ of $U$ there is a function $f \in \mathcal{F}$ that is injective on $U'$. Moreover, the family $\mathcal{F}$ has size $2^{O(k)} \log^2 n$ and any function evaluation on a given input takes $O(1)$ time. We use such a family $\mathcal{F}$ with $U = (B \setminus D) \cup C$ and $X$ being a set of size $O(k') = O(kc)$ where $k' = k(t - d) \leq k$. We refer to the elements of $X$ as colours.

Now for each function $f \in \mathcal{F}$, we will compute families $\mathcal{C}_1, \mathcal{C}_2, \ldots$, where for $j \geq 1$ each $\mathcal{C} \in \mathcal{C}_j$ is a sequence of $j$ pairwise disjoint $k$-subsets of $X$ that can be realised as the colour set via $f$ of a matching in $H$, i.e. there is an ordered matching $(e_1, \ldots, e_j)$ in $H$ such that $\mathcal{C} = (f(e_1), \ldots, f(e_j))$. Writing $\bigcup \mathcal{C} := \bigcup_{\mathcal{C} \in \mathcal{C}}$, we will construct $\mathcal{C}_j$ so that each $kj$-subset of $X$ that can be realised as the colour set of the vertices covered by a $j$-matching in $H$ occurs exactly once as $\bigcup \mathcal{C}$ for some $\mathcal{C} \in \mathcal{C}_j$.

To construct $\mathcal{C}_1$ we consider each $e \in H[U]$ sequentially and add $f(e)$ to $\mathcal{C}_1$ unless it is already present. Now suppose we have constructed $\mathcal{C}_j$ for some $j \geq 1$. Note that $|\mathcal{C}_j| \leq \binom{|X|}{kj}$, as each $kj$-subset of $X$ is represented at most once. To construct $\mathcal{C}_{j+1}$ we consider each $\mathcal{C} \in \mathcal{C}_j$ and $e \in H[U]$ sequentially, and add to $\mathcal{C}_{j+1}$ the sequence $\mathcal{C}'$ obtained by appending $f(e)$ to $\mathcal{C}$, unless $f(e) \cap \bigcup \mathcal{C} \neq \emptyset$ or $f(e) \cup \bigcup \mathcal{C}$ already occurs as $\bigcup \mathcal{C}''$ for some $\mathcal{C}'' \in \mathcal{C}_{j+1}$. We stop after we have constructed $\mathcal{C}_{t-d}$. The running time of the above process is

$$|\mathcal{F}| \cdot O((t - d) \cdot 2^{2|X|} \cdot |E(H)|) = 2^{O(kc)} \log^2 n \cdot O(c \cdot 2^{O(kc)} \cdot n^k) = O(n^k 2^{O(kc)}).$$

If the construction returns $\mathcal{C}_{t-d} = \emptyset$ for every $f \in \mathcal{F}$ then by the defining property of the perfect hash family $\mathcal{F}$ we conclude that $H[U]$ does not contain a matching of size $t - d$. Otherwise, suppose we obtain $\mathcal{C}^* \in \mathcal{C}_{t-d}$ for some $f \in \mathcal{F}$. Denote the $j$-th set in $\mathcal{C}^*$ by $\{x_{(j-1)k+1}, \ldots, x_{jk}\}$, for $j \in [t - d]$. By construction, $H[f^{-1}(x_{(j-1)k+1}), \ldots, f^{-1}(x_{jk})]$ induces at least one edge $e_j$ for each $j$, which can be found in time $O(|E(H)|) = O(n^k)$. The edges $e_1, \ldots, e_{t-d}$ form a matching of size $t - d$ in $H[U]$.

The overall running time is $O(n^k 2^{O(kc)})$.

5.5. Finding a cleaner. We conclude this section by proving the three lemmas that find a cleaner according to the various cases of the proof of Theorem 3.5.

Proof of Lemma 5.4. Assume that $H$ contains no even edges and $H \notin H_{n,k}$. Then the slack $s = |A| + |B| - n/k$ must be even. We apply Lemma 5.10 to find a matching $M$ covering $B$, which must consist of $BC^{k-1}$ edges, as there is no 2-edge. This preserves the slack, i.e. $s' = s = 0$, so $M$ is a cleaner, as required.

Proof of Lemma 5.5. We start with the case $s = 0$. Procedure PerfectMatchingEXT finds a matching $M_0$ of size $t - d = |B| - |D|$ in $(B \setminus D) \cup C$, as otherwise it would output “no PM”. By Lemma 5.11, we can enlarge $M_0$ to a matching $M$ of size $|B|$ covering $B$ and $(k - 1)|B|$ vertices in $C$. Thus we preserve the slack, i.e. $s' = s = 0$, so $M$ is a cleaner, as required.

It remains to consider $s > 0$. By assumption, $H$ has an $i$-edge for some even $i \in [2,k]$ and does not have any $ABC^{k-2}$ edge. Thus whenever we apply Lemma 5.10, we always obtain edges of the form $BC^{k-1}$. If $s$ is even then we can simply apply Lemma 5.10 to construct a matching $M$ covering $B$ with edges of the form $BC^{k-1}$. Indeed, this preserves the slack, i.e. $s' = s$, so $M$ is a cleaner, as required.

Finally, we can assume that $s$ is odd. By assumption, there is an $i$-edge $e_0$ with $i$ even. We fix any $e_0$ that minimises $i$. Note that $s \geq i-1$, otherwise Procedure PerfectMatchingEXT would output “no PM”. We apply Lemma 5.10 to construct a matching $M$ consisting of $e_0$ and $|B \setminus e_0|$ edges covering...
$B \setminus e_0$ with edges of the form $B C^{k-1}$. Thus we obtain slack $s' = |A \setminus V(M)| - (n/k - |M|) = s - i + 1$, which is even and non-negative, so again $M$ is a cleaner, as required.

**Proof of Lemma 5.6.** Suppose $s > 0$ and $H$ has some $A B C^{k-2}$ edge $e_0$, i.e. $|e_0 \cap A| = |e_0 \cap B| = 1$, and Procedure `PerfectMatchingEXT` outputs “PM”. We write $e_0 \cap B = \{x\}$. Our proof will use two different strategies for finding a cleaner that in combination cover all possible cases. The main strategy (which applies to all but one case) finds in time $O(n^k)$ a matching $M = M_1 \cup M_2$ covering $B$, such that, writing $s_2 := s_{M_2} = (|A \cup B| \setminus V(M_2)) - (n/k - |M_2|)$,

1. every edge $e$ of $M$ has $|e \cap B| = 1$ and $|e \cap A| \leq 1$,
2. $|M_1| \leq t + 1 = n/k - |A| + 1 = |B| - s + 1$,
3. the number of edges of $M_1$ that intersect $A$ is $(s_2 \mod 2) \in \{0, 1\}$. Note that if we find such a matching $M$ then $|M_2| = |M| - |M_1| \leq |B| - (t + 1) = s - 1$ and $s_2 = s + |M_2| - |(A \cup B) \setminus V(M_2)| \geq s - |M_2| \geq 1$, so $s_M = s_2 - (s_2 \mod 2)$ is even and non-negative, i.e. $M$ is a cleaner, as required. We will apply the main strategy to 3 cases below.

**Case 1.** Suppose $D = B$. By Lemma 5.11, we can find a matching $M'$ covering $B$ by $B C^{k-1}$ edges such that $V(M')$ is disjoint from $e_0 \setminus B$. Then $|M'| = |B| = t + s \geq t + 1$, as $s > 0$. Let $e_x$ be the edge of $M'$ containing $x$ (the vertex in $e_0 \cap B$). We let $M_2 = \emptyset$, so $s_2 = s = |A| + |B| - n/k$, and let $M_1 = M'$ if $s_2$ is even or otherwise let $M_1 = (M' \setminus \{e_x\}) \cup \{e_0\}$. Then $M$ satisfies (1-3).

**Case 2.** Suppose $t \leq |D| < |B|$. Fix any $v \in B \setminus D$. By Lemma 5.12, we can find an $A B C^{k-2}$ edge $e_v'$ containing $v$. Moreover, as $C$ is a maximal independent set, we can find a $B C^{k-1}$ edge $e_v$ containing $v$. Next, by Lemma 5.11 we can find a matching $M'$ covering $B$ by $B C^{k-1}$ edges with $V(M') \cap (\{e_v \cup e_v'\}) = \emptyset$. Then by Lemma 5.10 we find a matching $M_2$ covering $B \setminus (D \cup \{v\})$ with $V(M_2) \cap (\{e_v \cup e_v' \cup V(M')\}) = \emptyset$. We let $M_1 = M' \cup \{e_v\}$ if $s_2$ is even or otherwise let $M_1 = (M' \setminus \{e_v'\}) \cup \{e_v\}$. Then $M$ satisfies (1-3).

**Case 3.** Suppose $|D| = d < t$ and there is a $B C^{k-1}$ edge $e_*$ disjoint from $D \cup V(M_0)$, where $M_0$ is the matching of size $t - d$ in $(B \setminus D) \cup C$ found by Procedure `PerfectMatchingEXT`. We write $e_* \cap B = \{v\}$. By Lemma 5.11, we can find $B C^{k-1}$ edges covering $D$ that extend $M_0 \cup \{e_*\}$ to a matching $M'$ of size $t + 1$. Since $v \in D$, by Lemma 5.12 we have $\deg(v, A C^{k-2}) > k|B||A|^{(\frac{|C|}{k-3})} \geq k|M'||A|^{(\frac{|C|}{k})}$, so we can find an $A B C^{k-2}$ edge $e_v$ containing $v$ such that $e_v \cap V(M') = \{v\}$. By Lemma 5.10 we can find $M_2$ covering $B \setminus V(M')$ such that $V(M_2) \cap (V(M') \cup e_v) = \emptyset$. We let $M_1 = M'$ if $s_2$ is even or otherwise let $M_1 = (M' \setminus \{e_*\}) \cup \{e_v\}$. Then $M$ satisfies (1-3).

Finally, we describe the second strategy for finding a cleaner, which will complete the proof when none of the above 3 cases apply, i.e. when $|D| = d < t$ and there is no $B C^{k-1}$ edge $e_*$ disjoint from $D \cup V(M_0)$. By Lemma 5.11, we can find $B C^{k-1}$ edges covering $D$ that extend $M_0$ to a matching $M'$ of size $t$.

Then we claim that we can find $A B C^{k-2}$ edges covering $B \setminus V(M')$ that extend $M'$ to a matching $M$. To see this, we apply a greedy algorithm, where in each step, to cover some vertex $x$ of $B \setminus V(M')$ we consider any $B C^{k-2}$ set $S$ containing $x$ disjoint from all previous edges. By our assumptions for this case, $N(S)$ is disjoint from $C$, so as $|N(S)| \geq n/k - c > k|B|$ we can complete $S$ to an $A B C^{k-2}$ edge as required, so we can construct $M$ as claimed.

Writing $t' = |V(M_0) \cap B|$, we note that $M_0$ reduces the slack by $t' - |M_0| = t' - t + d$, and $M' \setminus M_0$ keeps it unchanged. Finally, $M' \setminus M_0$ reduces the slack by $|M' \setminus M| = |B| - d - t'$. Therefore $s_M = s - (t' - t + d) - (|B| - t' - d) = 0$, so $M$ is a cleaner. □

6. Concluding remarks

In this paper, we showed that $\text{DPM}_k(n, m)$ is in $P$ for $m \geq n/k - c$ for $c = O(\log n)$. For simplicity of analysis we did not attempt to optimise the exponent in the running time. We remark that some
improvements can be obtained by merging ‘transferrals’ and working with two partitions as in [14],
and/or settling for just testing for a perfect matching rather than actually finding one.

As observed in [14], the argument in [35] actually shows that \( \text{DPM}_k(n, n/k−n^\epsilon) \) is NP-complete
for any \( \epsilon > 0 \). Thus the complexity remains unknown for \( n/k−n^\epsilon \leq m < n/k−\omega(\log n) \). Our algorithm
is valid for any \( c = o(n) \), but our proof only gives polynomial running time when \( c = O(\log n) \). The
bottleneck to improving this comes from the extremal case, where we showed that the existence of
a perfect matching in \( H \) is equivalent to that of a matching of size \( t := n/k−|A| \) in \( H[B\cup C] \). Thus
we have a reduction to the following potentially simpler problem, which we believe has independent
interest.

Problem 6.1. Suppose \( 1/n \ll \epsilon \ll 1/k \). Let \( X, Y \) be disjoint sets with \( |X| = \epsilon n \) and \( |Y| = n \). Let
\( c = c(n) \) and \( t = t(n) \) with \( c \leq t \leq |X| \). Let \( H \) be a \( k \)-graph on \( X \cup Y \) with \( \delta_{k-1}(H) \geq t−c \) such that
\( H[Y] \) is independent. What is the complexity of deciding the existence of a matching of size \( t \) in \( H \)?

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