Approximation Algorithms for Multi-Dimensional Vector Assignment Problems

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Abstract

We consider a special class of axial multi-dimensional assignment problems called \textit{multi-dimensional vector assignment} (MVA) problems. An instance of the MVA problem is defined by \( m \) disjoint sets, each of which contains the same number \( n \) of \( p \)-dimensional vectors with nonnegative integral components, and a cost function defined on vectors. The cost of an \( m \)-tuple of vectors is defined as the cost of their component-wise maximum. The problem is now to partition the \( m \) sets of vectors into \( n m \)-tuples so that no two vectors from the same set are in the same \( m \)-tuple and so that the total cost of the \( m \)-tuples is minimized. The main motivation comes from a yield optimization problem in semi-conductor manufacturing. We consider two classes of polynomial-time heuristics for MVA, namely, \textit{hub} heuristics and \textit{sequential} heuristics, and we study their approximation ratio. In particular, we show that when the cost function is monotone and subadditive, hub heuristics, as well as sequential heuristics, have finite approximation ratio for every fixed \( m \). Moreover, we establish better approximation ratios for certain variants of hub heuristics and sequential heuristics when the cost function is monotone and submodular, or when it is additive. We provide examples to illustrate the tightness of our analysis. Furthermore, we show that the MVA problem is APX-hard even for the case \( m = 3 \) and for binary input vectors. Finally, we show that the problem can be solved in polynomial time in the special case of binary vectors with fixed dimension \( p \).

Keywords: multi-dimensional assignment; approximability; worst-case analysis; submodularity; wafer-to-wafer integration;

1. Introduction

1.1. Problem statement

We consider a multi-dimensional assignment problem motivated by an application arising in the semi-conductor industry. Formally, the input of the problem is defined by
m disjoint sets $V_1, \ldots, V_m$, where each set $V_k$ contains the same number $n$ of $p$-dimensional vectors with nonnegative integral components, and by a cost function $c(u) : \mathbb{Z}^p_+ \rightarrow \mathbb{R}_+$. Thus, the cost function assigns a nonnegative cost to each $p$-dimensional vector. A (feasible) $m$-tuple is an $m$-tuple of vectors $(u^1, u^2, \ldots, u^m) \in V_1 \times V_2 \times \ldots \times V_m$, and a feasible assignment for $V \equiv V_1 \times \ldots \times V_m$ is a set $A$ of $n$ feasible $m$-tuples such that each element of $V_1 \cup \ldots \cup V_m$ appears in exactly one $m$-tuple of $A$. We define the component-wise maximum operator $\lor$ as follows: for every pair of vectors $u, v \in \mathbb{Z}^p_+$,

$$u \lor v = (\max(u_1, v_1), \max(u_2, v_2), \ldots, \max(u_p, v_p)).$$

Now, the cost of an $m$-tuple $(u^1, \ldots, u^m)$ is defined as $c(u^1 \lor \ldots \lor u^m)$ and the cost of a feasible assignment $A$ is the sum of the costs of its $m$-tuples: $c(A) = \sum_{(u^1, \ldots, u^m) \in A} c(u^1 \lor \ldots \lor u^m)$.

With this terminology, the multi-dimensional vector assignment problem (MVA-$m$, or MVA for short) is to find a feasible assignment for $V$ with minimum total cost. A case of special interest is the case when all vectors in $V_1 \cup \ldots \cup V_m$ are binary 0–1 vectors; we call this special case binary MVA. Finally, the wafer-to-wafer integration problem (WWI-$m$ or WWI for short) arises when the cost function of the binary MVA is additive, meaning that $c(u) = \sum_{i=1}^p u_i$.

In this paper, we investigate how closely the optimal solution of MVA-$m$ and WWI-$m$ can be approximated by polynomial-time approximation algorithms.

**Example 1.** An instance of WWI with $m = 3, n = p = 2$ is displayed in Figure 1.1. The optimal value of the instance is equal to 2: it is achieved by assigning the first vector of $V_1$, the second vector of $V_2$, and the first vector of $V_3$ to the same triple, thus arriving at vector $(1, 0)$ with cost $c(1, 0) = 1$; the remaining three vectors form a second triple with cost $c(0, 1) = 1$.

![Figure 1: A WWI-3 instance with $m = 3, n = p = 2$](image)

1.2. Wafer-to-wafer integration and related work

The motivation for studying the WWI problem arises from the optimization of the wafer-to-wafer production process in the electronics industry. We only provide a brief description of this application; for additional details, we refer to papers by Reda, Smith and Smith [11], Taouil and Hamdioni [17], Taouil et al. [18], and Verbree et al. [19].

For our purpose, a wafer can be viewed as a string of elements called dies. Each die can be either good (operative) or bad (defective). So, a wafer can be modeled as a binary
vector, where each ‘0’ represents a good die and each ‘1’ represents a bad die. There are $m$ lots of wafers, say $V_1, \ldots, V_m$, and each lot contains $n$ wafers. All wafers in a given lot are meant to have identical functionalities, were it not for the occasional occurrence of defective dies during the previous production steps. The wafer-to-wafer integration process requires to form stacks, where a stack is obtained by “superposing” $m$ wafers chosen from different lots; thus, a stack corresponds to a feasible $m$-tuple. As a result of integration, each position in the stack gives rise to a three-dimensional stacked integrated circuit (3D-SIC) which is ‘good’ only when the corresponding $m$ entries of the selected wafers are ‘good’; otherwise, the 3D-SIC is ‘bad’. The yield optimization problem now consists in assigning the available wafers to $n$ stacks so as to minimize the total number of bad 3D-SICs. Thus, the WWI problem provides a model for yield optimization.

The wafer-to-wafer yield optimization problem has recently been the subject of much attention in the engineering literature. One example is the contribution by Reda et al. [11]. These authors formulate WWI as a multi-dimensional assignment problem. A natural formulation of WWI as an integer linear programming problem turns out to be hard to solve to optimality for instances with large values of $m$ (typical dimensions for the instances are: $3 \leq m \leq 10$, $25 \leq n \leq 75$, $500 \leq p \leq 1000$). On the other hand, Reda et al. [11] propose several heuristics and show that they perform well in computational experiments. Some recent work in this direction is also reported in [13, 17, 18, 19].

Our main objective in this paper is to study the approximability of the MVA problem and of the WWI problem (in the sense of [20]). Let us note at this point that the wafer-to-wafer integration problem is usually formulated in the literature as a maximization problem (since one wants to maximize the yield). However, we feel that from the approximation point of view, it is more appropriate to study its cost minimization version. Indeed, in industrial instances, the number of bad dies in each wafer is typically much less than the number of good dies. Therefore, it is more relevant to be able to approximate the (smaller) minimum cost than the (larger) maximum yield.

Since MVA is defined as a multi-dimensional assignment problem with a special cost structure, our work relates to previous publications on special classes of multi-dimensional assignment problems, such as Bandelt, Crama and Spieksma [1], Burkard, Rudolf and Woeginger [3], Crama and Spieksma [4], Dokka, Kouvela and Spieksma [6], Goossens et al. [7], Spieksma and Woeginger [15], etc. Surveys on multi-dimensional assignment problems can be found in Chapter 10 of [2] and in [14]. To the best of our knowledge, the approximability of MVA has only been previously investigated by Dokka et al. [5], who mostly focused on the case $m = 3$ with additive cost functions. The present paper extends to MVA-$m$ and considerably strengthens the results presented in [5].

1.3. Contents of the paper

Section 2 contains a formulation of the problem as an integer program (Subsection 2.1), discusses various possible assumptions on the cost function (Subsection 2.2), describes two classes of heuristics (Subsection 2.3), and gives an overview of our results in Subsection 2.4. In Section 3, we prove that the heuristics have finite worst-case performance for every fixed $m$ under various assumptions on the cost function $c$. In Section 4, we prove that the WWI-$m$ problem is APX-hard even when $m = 3$, all input vectors are binary, and the cost function is additive. Finally, we show in Section 5 that WWI-$m$ can be solved in polynomial time when $p$ is fixed.
2. Problem formulation, properties, heuristics and results

2.1. Problem formulation

As mentioned before, let \( V = V_1 \times V_2 \times \ldots \times V_m \) be the set of feasible \( m \)-tuples. By a slight abuse of notations, we write \( u^k \in a \) when \( a = (u^1, \ldots, u^m) \) and \( 1 \leq k \leq m \). We also extend the definition of \( c \) to \( m \)-tuples of \( \mathbb{Z}_{+}^m \) by setting \( c(u^1, \ldots, u^m) := c(u^1 \lor \ldots \lor u^m) \), and, when \( W \) is any set of \( m \)-tuples, we define \( c(W) = \sum_{a \in W} c(a) \).

Let us first provide an IP-formulation of MVA-\( m \) as an \( m \)-dimensional axial assignment problem. For each \( a \in V \), let \( x_a \) be a binary variable indicating whether \( m \)-tuple \( a \) is selected (\( x_a = 1 \)) or not (\( x_a = 0 \)) in the optimal assignment. Reda et al. [11] give the following formulation of WWI, which directly extends to MVA:

\[
\begin{align*}
\text{minimize} & \quad \sum_{a \in V} c(a)x_a \\
\text{s.t.} & \quad \sum_{u^k \in a} x_a = 1 \quad \text{for all } u \in \bigcup_{i=1}^{m} V_i, \\
& \quad x_a \in \{0, 1\} \quad \text{for all } a \in V.
\end{align*}
\]

Other formulations of MVA exist; for instance, Dokka et al. [5] propose an alternative IP-formulation that may be more effective from a computational perspective.

In any application of MVA, the cost function \( c \) is likely to have some structure. Indeed, in the WWI-application motivating this study, we have, as mentioned before, an additive cost function: \( c(u) = \sum_{i=1}^{p} u_i \). We now list various possible assumptions on the cost function \( c \).

2.2. Properties of the cost function \( c \)

We focus our attention on cost functions \( c(u) \) satisfying one or more of the following properties:

**Monotonicity:** The cost function \( c \) is monotone if, for all \( u, v \in \mathbb{Z}_{+}^p \) with \( u \leq v \), we have \( 0 \leq c(u) \leq c(v) \).

**Subadditivity:** The cost function \( c \) is subadditive if, for all \( u, v \in \mathbb{Z}_{+}^p \), we have \( c(u \lor v) \leq c(u) + c(v) \).

**Submodularity:** The cost function \( c \) is submodular if, for all \( u, v \in \mathbb{Z}_{+}^p \), we have \( c(u \lor v) + c(u \land v) \leq c(u) + c(v) \).

(Here, \( \land \) denotes the component-wise minimum operator:
\[
u \land v = (\min(u_1, v_1), \min(u_2, v_2), \ldots, \min(u_p, v_p)).
\]

Submodular cost functions frequently appear in the analysis of approximation algorithms for combinatorial optimization problems; for recent illustrations, see for instance [9, 16] and the references therein. The additive cost function of problem WWI actually satisfies a much stronger property than submodularity, namely:

**Modularity:** The cost function \( c \) is modular if, for all \( u, v \in \mathbb{Z}_{+}^p \), we have \( c(u \lor v) + c(u \land v) = c(u) + c(v) \).
It is well-known that \( c \) is modular if and only if there exist \( p \) functions \( f_\ell(u_\ell) \) such that \( c(u) = \sum_{\ell=1}^{p} f_\ell(u_\ell) \) (see, e.g., Theorem 2.3.3 in Simchi-Levy, Chen and Bramel [12]). For the MVA problem, therefore, assuming additivity is essentially equivalent to assuming monotonicity and modularity.

2.3. Heuristics

Consider any heuristic algorithm \( H \) for MVA-\( m \). Following standard terminology (see, e.g., Williamson and Shmoys [20]), we say that \( H \) is a \( \rho^H(m) \)-approximation algorithm for MVA-\( m \) if \( H \) runs in polynomial time and if \( \rho^H(m) \) is (an upper bound on) the approximation ratio of \( H \), in the following sense: for every instance of MVA-\( m \) with optimal value \( c_{\text{OPT}}^m \), when \( H \) returns the assignment \( A_m \), then \( c(A_m) \leq \rho^H(m)c_{\text{OPT}}^m \).

Here, we are interested in the behavior of the following hub and sequential heuristics, which all rely on the observation that MVA-2 boils down to a classical bipartite assignment (or matching) problem (see, e.g., Bandelt et al. [1] for other examples of sequential and hub heuristics).

We first describe so-called hub heuristics, where one particular set \( V_h \) acts as a “hub”, and where a feasible solution is obtained by combining bipartite assignments constructed for each pair \((V_h, V_i)\) into a feasible assignment; see Algorithm 1. We will also analyze a version of the single-hub heuristic called heaviest-hub heuristic, or \( H_{\text{hhub}} \); here, the hub \( V_h \) is the heaviest set, that is, \( c(V_h) \geq c(V_k) \) for \( k = 1, \ldots, m \); see Algorithm 2. The idea underlying this initial condition is to make sure that all assignments will be able to take the “worst lot” into account.

Finally, since there is one feasible single-hub solution for each possible choice of the hub \( V_h, \ h = 1, \ldots, m \), we call multi-hub heuristic the heuristic that outputs the best of these \( m \) solutions; see Algorithm 3.

Algorithm 1 Single-hub heuristic \( H_{\text{hub}}(V_h) \)

\[
\text{function } H_{\text{hub}}(V_h) \\
\text{\{comment: } h \in \{1,\ldots,m\} \text{ is the index of the hub\}} \\
\text{for } i = 1 \text{ to } m, i \neq h, \text{ do} \\
\text{\quad solve an assignment problem between } V_h \text{ and } V_i, i \neq h, \text{ based on costs } c(u \lor v), u \in V_h, v \in V_i; \text{ call the resulting optimal assignment } M_{hi}; \text{ say, } M_{hi} = \{(u^1_i, u^2_i) \mid u^1_i \in V_h, u^2_i \in V_i, j = 1, \ldots, n\}; \\
\text{end for} \\
\text{construct the feasible solution } M_h = \{(u^1_j, u^2_j, \ldots, u^m_j) \mid (u^1_j, u^2_j, \ldots, u^m_j) \in M_{hi}, i = 1, 2, \ldots, m, j = 1, \ldots, n\}; \\
\text{output } M_h.
\]

Algorithm 2 Heaviest-hub heuristic \( H_{\text{hhub}} \)

\[
\text{reindex } V_1, \ldots, V_m \text{ so that } c(V_1) \geq c(V_k) \text{ for } k = 1, \ldots, m; \\
\text{apply the single-hub heuristic } H_{\text{hub}}(V_1).
\]

Let us now turn to the sequential heuristic \( H_{\text{seq}} \) described as Algorithm 4: \( H_{\text{seq}} \) progressively builds a feasible solution \( H_m \) by optimally assigning the next set \( V_i \) to a partial solution \( H_{i-1} \). We point out that, for WWI-\( m \), Reda et al. [11] proposed an
Algorithm 3 Multi-hub heuristic $H^{mhub}$

for $h = 1$ to $m$

apply the single-hub heuristic $H^{hub}(V_h)$ to produce the feasible solution $M_h$;

end for

let $h^* = \arg\min_h c(M_h)$; output $M_{h^*}$.

Algorithm 4 Sequential heuristic $H^{seq}$

let $H_1 := V_1$;

for $i = 2$ to $m$

solve a bipartite assignment problem between $H_{i-1}$ and $V_i$ based on the costs $c(u^1 \lor \ldots \lor u^{i-1} \lor v)$, for all $(u^1, \ldots, u^{i-1}) \in H_{i-1}$ and $v \in V_i$; let $H_i$ be the resulting assignment for $V_1 \times V_2 \times \ldots \times V_i$;

end for

output $H_m$.

Algorithm 5 Heaviest-first heuristic $H^{heavy}$

reindex $V_1, \ldots, V_m$ so that $\max(c(V_1), c(V_2)) \geq c(V_k)$ for $k = 1, \ldots, m$;

apply the sequential heuristic.

Clearly, each of the above heuristics runs in polynomial time. In fact, one can measure the complexity of these heuristics by observing how many (two-dimensional) assignment problems they need to solve. The most expensive one is $H^{mhub}$, since it solves $O(m^2)$ assignment subproblems, whereas $H^{hub}$, $H^{hhub}$, $H^{seq}$, and $H^{heavy}$ only solve $O(m)$ subproblems. Observe that the preprocessing step needed for $H^{hhub}$ and $H^{heavy}$ does not increase their complexity.

2.4. Overview of results

In this section we list the main results proved in our paper. First, in case, $c$ is monotone and subadditive, no feasible solution can be arbitrarily far away from the optimum, as expressed by the next theorem.

Theorem 2. Every heuristic $H$ that returns a feasible solution is an $m$-approximation algorithm when the cost function $c$ is monotone and subadditive. The approximation ratio $\rho^H(m) = m$ is tight for all $m \geq 2$, even for WWI-$m$.
We prove this result in Subsection 3.1. Next, we establish that the multi-hub heuristic has an approximation ratio of \( \frac{m}{2} \) when \( c \) is monotone and submodular. In fact, this ratio already holds for the single-hub heuristic \( H_{\text{hub}}(V_h) \) when we assume that \( V_h \) is the heaviest set.

**Theorem 3.** The heaviest-hub heuristic \( H_{\text{hhub}} \) and the multi-hub heuristic \( H_{\text{mhub}} \) are \( \frac{m}{2} \)-approximation algorithms for MVA-\( m \) when the cost function \( c \) is monotone and submodular. The approximation ratio \( \rho_{\text{hhub}}(m) = \rho_{\text{mhub}}(m) = \frac{m}{2} \) is tight for all \( m \geq 2 \), even for binary MVA.

We prove this result in Subsection 3.2. Also, when \( c \) is monotone and submodular, the sequential heuristic has the same worst-case approximation ratio:

**Theorem 4.** The sequential heuristic \( H_{\text{seq}} \) is an \( \frac{m}{2} \)-approximation algorithm for MVA-\( m \) when the cost function \( c \) is monotone and submodular, for every order of the sets \( V_1, \ldots, V_m \). The approximation ratio \( \rho_{\text{seq}}(m) = \frac{m}{2} \) is tight for all \( m \geq 2 \), even for the heaviest-first heuristic and even for binary MVA.

We prove this result in Subsection 3.3. When \( c \) is additive, a better bound can be proved for the heaviest-first heuristic:

**Theorem 5.** The heaviest-first heuristic \( H_{\text{heavy}} \) is a \( \left(\frac{1}{2}(m+1) - \frac{1}{4}\ln(m-1)\right) \)-approximation algorithm for MVA-\( m \) when the cost function \( c \) is additive.

We prove this result in Subsection 3.4. Although we do not know whether the bound in Theorem 5 is tight, we exhibit in Section 3.5 a family of instances for which \( H_{\text{heavy}} \) displays the following behavior:

**Theorem 6.** There exists an infinite sequence of values of \( m \) such that the heaviest-first heuristic produces a feasible assignment with cost larger than \( \sqrt{\frac{m}{2}} c_{\text{OPT}} \) on certain instances of WWI-\( m \).

This concludes the overview of our results concerning approximation ratios of the heuristics; see also Figure 2.

One might wonder about the precise complexity status of MVA-\( m \), and of its special case WWI-\( m \). The following result implies that, when restricting ourselves to polynomial-time algorithms, constant-factor approximation algorithms are the best we can hope for (unless P=NP), even for WWI-3:

**Theorem 7.** WWI-3 is APX-hard, even when all vectors in \( V_1 \cup V_2 \cup V_3 \) are \( 0-1 \) vectors with exactly two nonzero entries per vector.

We prove this in Section 4. Finally, in case the dimension \( p \) of the vectors is fixed, we show in Section 5 that binary MVA-\( m \) can be solved in polynomial time:

**Theorem 8.** Binary MVA can be solved in polynomial time for each fixed \( p \).

### 3. Proofs of approximation ratios

This section is devoted to the proofs of the approximation ratios of the heuristics described in Subsection 2.3.
3.1. Monotone and subadditive costs: feasible solutions

Here, we first establish some properties of feasible solutions depending on various assumptions on the cost function \( c \). Consider a feasible assignment \( A_m \) for \( V_1 \times \ldots \times V_m \), and let \( A_k \) denote the restriction of this assignment to \( V_1 \times \ldots \times V_k \), for all \( k \leq m \). Denote by \( c^{OPT}_k \) the optimal value of the restricted instance \( V_1 \times \ldots \times V_k \).

**Lemma 9.** If the cost function \( c \) is monotone and if \( A_m \) is a feasible assignment, then, for all \( i \leq k \leq m \),
\[
c(V_i) \leq c^{OPT}_k \leq c(A_k) \leq c(A_m).
\] (3.1)

**Proof.** Obvious. □

**Lemma 10.** If the cost function \( c \) is subadditive and if \( A_m \) is a feasible assignment, then
\[
c(A_m) \leq c(A_{m-1}) + c(V_m).
\] (3.2)

**Proof.** Assume without loss of generality that the \( j^{th} \) \( m \)-tuple of \( A_m \) is \((u^1_j, \ldots, u^m_j)\) (that is, the \( j^{th} \) \( m \)-tuple in the assignment contains the \( j^{th} \) vector of \( V_i \) for each \( i \)). Then,
\[
c(A_m) = \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^m_j)
\leq \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^{m-1}_j) + \sum_{j=1}^{n} c(u^m_j)
= c(A_{m-1}) + c(V_m).
\]

These two lemmas allow us to prove:
Theorem 2. Every heuristic $H$ that returns a feasible solution is an $m$-approximation algorithm when the cost function is monotone and subadditive. The approximation ratio $\rho_H(m) = m$ is tight for all $m \geq 2$, even for WWI-$m$.

Proof. The statement holds for $m = 1$. Then, using Eq. (3.2) from Lemma 10, Eq. (3.1) from Lemma 9, and induction on $m$, we obtain that every feasible solution $A_m$ satisfies

$$c(A_m) \leq c(A_{m-1}) + c(V_m) \leq (m - 1)c_{OPT}^{m-1} + c_{OPT}^m$$

To see that the bound is tight, let $p = 1$, $n = m$, $V_i = \{1, 0, \ldots, 0\}$ for $i = 1, \ldots, m$, and $c(u) = u$ for all $u \in \mathbb{R}$. The cost function is obviously additive, hence this is an instance of WWI-$m$. The worst feasible assignment yields $\{1,1,\ldots,1\}$ with cost $m$, whereas the optimal assignment has cost 1.

Thus, Theorem 2 implies that every heuristic has bounded worst-case performance (for fixed $m$) under the assumption that $c$ is monotone and subadditive. On the other hand, if we relax either of the assumptions on $c$, then even the heaviest-hub and heaviest-first sequential heuristics do not have bounded approximation ratios on WWI-3, as shown by the following examples.

Example 11. For any $p$, we denote by $0$, $1$, and $e_i$, respectively, the all-zero, all-one, and $i$-th unit vector of $\mathbb{Z}^p$.

Let $p = 3$, $V_1 = \{e_1, 0\}$, $V_2 = \{0, e_2\}$, $V_3 = \{1, 0\}$, and $c(u) = u_1 + u_2 + u_3 - 3\min(u_1, u_2, u_3)$. This cost function is nonnegative, subadditive (and even submodular), but not monotone since $1 = c(e_1) \not\leq c(1) = 0$, while $e_1 \leq 1$. The optimal solution for this instance is $\{(e_1, e_2, 1), (0, 0, 0)\}$ with cost 0. Since $c(V_1) = c(V_2) > c(V_3)$, the heaviest-first heuristic could match $V_1$, $V_2$ to produce $\{(e_1, 0), (0, e_2)\}$, then $V_3$ to produce $\{(e_1, 0, 1), (0, e_2, 0)\}$ with cost 1. Heaviest-hub can produce the same solution.

A similar observation applies when $c$ is not subadditive: let $p = 3$, $V_1 = \{e_1, 0\}$, $V_2 = \{0, e_2\}$, $V_3 = \{e_3, e_3\}$, and $c(u) = u_1 + u_2 + M\min(u_1, u_2, u_3)$ with $M > 0$. This cost function is nonnegative, monotone (and supermodular), but not subadditive since $c(1, 0, 1) = c(0, 1, 1) = 1$ and $c(1, 1, 1) = 2 + M$. The optimal solution is $\{(e_1, 0, e_3), (0, e_2, e_3)\}$ with cost 2. Note that $c(V_1) = c(V_2) = 1$, $c(V_3) = 0$; hence, heaviest-first could match $V_1$, $V_2$ to produce $\{(e_1, e_2, 1), (0, 0)\}$, then $V_3$ to produce $\{(e_1, e_2, 3), (0, 0, e_3)\}$ with cost $M + 2$. So, the performance of heaviest-first (and similarly, heaviest-hub) is unbounded for this instance.

3.2. Monotone and submodular costs: hub heuristics

The proof of Theorem 2 easily implies that the ratio $\rho_{hub}(m) = m - 1$ is valid for the solution produced by any single-hub heuristic when the cost function is monotone and subadditive (simply start the induction with $m = 2$ in the proof). This ratio is actually tight: To see this, consider an arbitrary instance $I$ of MVA-$(m - 1)$, and extend it with an additional set $V_m$ consisting of $n$ zero vectors. With $V_m$ as the hub, $H^{hub}(V_m)$ can produce any feasible solution of $I$. Hence, Theorem 2 establishes the tightness of the bound.
We are going to show next that, for heaviest-hub and multi-hub heuristics, better approximation ratios can be established when we assume that the cost function is monotone and submodular.

In the sequel, we frequently assume without loss of generality, as in the proof of Lemma 10, that the $j^{th}$ $m$-tuple of $A_m$ is $(u^1_j, \ldots, u^m_j)$. Under this assumption, we now derive inequalities that are valid for every feasible assignment $A_m$.

**Lemma 12.** If the cost function $c$ is monotone and submodular, and if $A_m$ is a feasible assignment such that the $j^{th}$ $m$-tuple of $A_m$ is $(u^1_j, \ldots, u^m_j)$, then, for all $k \in \{1, \ldots, m-1\}$,

\[
c(A_m) \leq c(A_{m-1}) + \sum_{j=1}^{n} c(u^k_j \lor u^m_j) - c(V_k) \leq c(A_{m-1}) + c(V_m) - \sum_{j=1}^{n} c(u^k_j \land u^m_j) \leq c(A_{m-1}) + c(V_m).
\]

**Proof.**

\[
c(A_m) = \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^m_j) \leq \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^{m-1}_j) + \sum_{j=1}^{n} c(u^k_j \lor u^m_j) - \sum_{j=1}^{n} c((u^1_j \lor \ldots \lor u^{m-1}_j) \land (u^k_j \lor u^m_j)) \leq \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^{m-1}_j) + \sum_{j=1}^{n} c(u^k_j \lor u^m_j) - \sum_{j=1}^{n} c(u^k_j) \leq \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^{m-1}_j) + \sum_{j=1}^{n} c(u^m_j) - \sum_{j=1}^{n} c(u^k_j \land u^m_j) \leq \sum_{j=1}^{n} c(u^1_j \lor \ldots \lor u^{m-1}_j) + \sum_{j=1}^{n} c(u^m_j)
\]

where (3.6) is by definition of the cost function, (3.7) holds by submodularity applied to $u = u^1_j \lor \ldots \lor u^{m-1}_j$ and $v = u^k_j \lor u^m_j$ for each $j$, (3.8) follows by monotonicity (since $u^k_j \leq (u^1_j \lor \ldots \lor u^{m-1}_j) \land (u^k_j \lor u^m_j)$), (3.9) by submodularity applied to $u = u^k_j$, $v = u^m_j$, and (3.10) by nonnegativity of $c$. Inequalities (3.8), (3.9), (3.10) are equivalent to (3.3), (3.4), (3.5), respectively.

We can now prove:

**Theorem 3.** The heaviest-hub heuristic $H^{hhub}$ and the multi-hub heuristic $H^{mhub}$ are $\frac{m}{2}$-approximation algorithms for MVA-$m$ when the cost function $c$ is monotone and submodular. The approximation ratio $\rho^{hhub}(m) = \rho^{mhub}(m) = \frac{m}{2}$ is tight for all $m \geq 2$. 

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even for binary MVA.

**Proof.** We prove the theorem by induction on $m$. The result is trivial when $m = 2$. For larger values of $m$, assume as in the description of Algorithm 2 that $V_1$ is the heaviest set, let $H_m = M_1$ be the solution found by the heaviest-hub heuristic $H^{hhub}$, and let $H_{m-1}$ be the restriction of this assignment $H_m$ to $W = V_1 \times \ldots \times V_{m-1}$.

We consider now two cases. Assume first that $c(V_1) \leq \frac{1}{2} c^{OPT}$. Applying (3.5) to the assignment $M_1$, we obtain

$$c(H_m) \leq c(H_{m-1}) + c(V_m).$$

(3.11)

Since $H_{m-1}$ results from applying the heaviest-hub heuristic (with heaviest hub $V_1$) to $W = V_1 \times \ldots \times V_{m-1}$, we have by induction and by monotonicity of $c$:

$$c(H_{m-1}) \leq \rho^{hhub}(m-1) c^{OPT}(W) \leq \frac{1}{2} (m-1) c^{OPT}_m$$

(3.12)

Finally, using the assumption that $c(V_1) \leq \frac{1}{2} c^{OPT}_m$, we conclude from (3.11)–(3.12) that

$$c(H_m) \leq \left( \frac{m-1}{2} + \frac{1}{2} \right) c^{OPT}_m = \frac{m}{2} c^{OPT}_m.$$

Consider next the case where $c(V_1) > \frac{1}{2} c^{OPT}_m$. Assume, without loss of generality, that the $j^{th}$ component of $H_m$ is $(u^1_j, \ldots, u^m_j)$. Then, by Eq. (3.3):

$$c(H_m) \leq c(H_{m-1}) + \sum_{j=1}^{n} c(u^1_j \lor u^m_j) - c(V_1).$$

With $M_{1,m}$ denoting the optimal matching of $V_1$ and $V_m$ as in Algorithm 1, we find: 

$$\sum_{j=1}^{n} c(u^1_j \lor u^m_j) = c(M_{1,m}) \leq c^{OPT}_m.$$ Thus,

$$c(H_m) \leq c(H_{m-1}) + c(M_{1,m}) - c(V_1)$$

$$\leq \rho^{hhub}(m-1) c^{OPT}_{m-1} + c^{OPT}_m - \frac{1}{2} c^{OPT}_m$$

$$\leq \left( \frac{m-1}{2} + \frac{1}{2} \right) c^{OPT}_m$$

$$= \frac{m}{2} c^{OPT}_m.$$

This proves that the approximation ratio $\rho^{hhub}(m) = \frac{1}{2} m$ is valid for the heaviest-hub heuristic $H^{hhub}$ and hence, for the multi-hub heuristic $H^{hhub}$ as well.

To prove that the ratio is tight, consider the function $r_2(u) = f(\sum_{i=1}^{p} u_i)$, where $f: \mathbb{R} \rightarrow \mathbb{R}$ is defined by $f(x) = x$ when $x \leq 2$, and $f(x) = 2$ when $x \geq 2$. Since $f$ is monotone nondecreasing and concave, it follows easily that $r_2$ is monotone and submodular on $\mathbb{Z}_+^p$ (see, e.g., Theorem 2.3.6 in Simchi-Levy et al. [12]). (When $u$ is a binary vector, $r_2(u)$ is the rank function of the uniform matroid of rank 2.)

Now, let $p = n = m$, $V_i = \{e_i, \mathbf{0}, \ldots, \mathbf{0}\}$ for $i = 1, \ldots, m$, and $c(u) = r_2(u)$. By symmetry, any of the sets $V_i$ can be chosen as the heaviest set, and the multi-hub heuristic delivers a solution with the same cost as the heaviest-hub heuristic. In particular, it is
easy to see that multi-hub can produce the assignment \( H_m \) in which \( e_i \) is matched with \( m-1 \) zero vectors, for all \( i \). The resulting assignment \( H_m \) has cost \( m \), whereas the optimal solution assigns \( (e_1, \ldots, e_m) \) to the same tuple, and has cost \( r_2(e_1 \lor \ldots \lor e_m) = 2 \). ■

Let us observe that the submodularity assumption is necessary in Theorem 3, as shown by the following example.

Example 13. Let \( m = 3 \), \( n = 2 \), \( p = 3 \), \( V_1 = \{e_1, e_1\} \), \( V_2 = V_3 = \{e_2, e_3\} \), and \( c(u) = \max(u_1, u_2, u_3) + \min(u_1, u_2, u_3) \). This cost function can be checked to be subadditive, but not submodular. The optimal solution is \( \{(e_1, e_2, e_2), (e_1, e_3, e_3)\} \), with cost \( 2 \). However, using \( V_1 \) as a hub, heaviest-hub may find the solution \( \{(e_1, e_2, e_2), (e_1, e_3, e_2)\} \) with cost \( 4 > \frac{m}{2} \). Multi-hub may fail in the same way.

### 3.3. Monotone and submodular costs: sequential heuristics

Let us now turn to the analysis of sequential heuristics. It follows again from the proof of Theorem 2 that the performance ratio of any sequential heuristics is bounded by \( m-1 \) when the cost function is monotone and subadditive. Under the stronger submodularity assumption, we can establish a better bound:

**Theorem 4.** The sequential heuristic \( H^{seq} \) is an \( \frac{m}{2} \)-approximation algorithm for MVA-\( m \) when the cost function \( c \) is monotone and submodular, for every order of the sets \( V_1, \ldots, V_m \). The approximation ratio \( \rho^{seq}(m) = \frac{m}{2} \) is tight for all \( m \geq 2 \), even for the heaviest-first heuristic and even for binary MVA.

**Proof.** Let \( H_m \) be a feasible assignment for \( V \) found by the sequential heuristic. We prove the theorem by induction on \( m \). The result is trivial when \( m = 2 \). For larger values of \( m \), we distinguish among two cases as in the proof of the previous theorem. Assume first that \( c(V_{m-1}) \leq \frac{1}{2}c^{OPT} \). Then, consider the partial assignment \( A_{m-2,m} \) that is obtained by assigning optimally \( V_m \) to \( H_{m-2} \) (independently of \( V_{m-1} \)). Let \( H_m^* \) be the concatenation of \( H_{m-1} \) and \( A_{m-2,m} \) (that is, \( H_m^* \) assigns \( V_{m-1} \) to \( H_{m-2} \) as in \( H_{m-1} \), and \( V_m \) to \( H_{m-2} \) as in \( A_{m-2,m} \)). Note that \( H_m^* = H_m \); therefore, \( c(H_m) \leq c(H_m^*) \) since, by definition, the sequential heuristic assigns \( V_m \) optimally to \( H_m \). Applying (3.5) to the assignment \( H_m^* \), we obtain

\[
c(H_m) \leq c(H_m^*) \leq c(A_{m-2,m}) + c(V_{m-1}). \tag{3.13}
\]

Since \( A_{m-2,m} \) results from applying the sequential heuristic to \( W = V_1 \times \ldots \times V_{m-2} \times V_m \), we have by induction:

\[
c(A_{m-2,m}) \leq \rho^{seq}(m-1)c^{OPT}(W) \leq \frac{1}{2}(m-1)c^{OPT}_m \tag{3.14}
\]

where \( c^{OPT}(W) \) is the cost of an optimal assignment on \( W \).

Finally, using the assumption that \( c(V_{m-1}) \leq \frac{1}{2}c^{OPT}_m \), we conclude from (3.13)–(3.14) that

\[
c(H_m) \leq \left( \frac{m-1}{2} + \frac{1}{2} \right)c^{OPT}_m = \frac{m}{2}c^{OPT}_m.
\]

Assume now alternatively that \( c(V_{m-1}) \geq \frac{1}{2}c^{OPT}_m \). Let \( M_{m-1,m} \) be an optimal matching of \( V_{m-1} \) with \( V_m \), and consider the assignment \( H_m^* \) obtained by concatenating \( H_{m-1} \) with \( M_{m-1,m} \). Assume, without loss of generality, that the \( j^{th} \) vector of \( H_m^* \) is \((u_1^j, \ldots, u_m^j)\).
Then, by definition of $H_m$, $c(H_m) \leq c(H_m^+)$ and by Eq. (3.3):

$$c(H_m) \leq c(H_m^+) \leq c(H_{m-1}) + \sum_{j=1}^{n} c(u_j^{m-1} \vee u_j^m) - c(V_{m-1}).$$

(3.15)

Moreover, $\sum_{j=1}^{n} c(u_j^{m-1} \vee u_j^m) = c(M_{m-1,m}) \leq c_{m}^{OPT}$. Thus, we derive

$$c(H_m) \leq c(H_{m-1}) + c(M_{m-1,m}) - c(V_{m-1})$$

$$\leq \rho_{seq}(m-1)c_{m-1}^{OPT} + c_{m}^{OPT} - \frac{1}{2}c_{m}^{OPT}$$

$$\leq \left( \frac{m-1}{2} + \frac{1}{2} \right)c_{m}^{OPT}$$

$$= \frac{m}{2}c_{m}^{OPT}.$$  

This establishes the validity of the approximation ratio $\rho_{seq}(m) = \frac{1}{2}m$. 

The example given in the proof of Theorem 3 proves that the approximation ratio $\rho_{seq}(m) = \frac{m}{2}$ is tight even for the heaviest-first heuristic. □

As a side-remark, the worst-case example used in the proof of Theorem 3 and of Theorem 4 shows that, for monotone submodular instances of binary MVA-$m$, the same ratio $\frac{m}{2}$ is tight for the (expensive) combined heuristic that results by successively running the single-hub heuristic and the sequential heuristic for all possible choices of the hub and for all possible permutations of the sets $V_1, \ldots, V_m$. Also, Example 13 shows that the submodularity assumption is necessary in Theorem 4.

We return in Section 3.5 to a discussion of the approximation ratio of sequential heuristics for the more restrictive WWI-$m$ problem.

3.4. Additive costs: heaviest-first heuristic

In this section, we explicitly rely on the assumption that the cost function is additive, i.e., $c(u) = \sum_{\ell=1}^{p} u_\ell$, and we derive an improved approximation ratio for the heaviest-first heuristic. We first establish a series of preliminary results.

3.4.1. Preliminary results for additive cost functions

If the $j^{th}$ $m$-tuple of an arbitrary assignment $A_m$ is $u_j = (u_j^1, \ldots, u_j^m)$, then, for all $j = 1, \ldots, n$

$$c(u_j) = \sum_{\ell=1}^{p} (u_j^1 \vee \ldots \vee u_j^\ell).$$

Thus,

$$c(A_m) = \sum_{j=1}^{n} c(u_j)$$

$$= c(A_{m-1}) + c(V_m) - \sum_{j=1}^{n} \sum_{\ell=1}^{p} (u_j^1 \vee \ldots \vee u_j^{m-1} \wedge u_j^\ell).$$

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For each \( j, \ell \), let \( k(j, \ell) \) be an (arbitrary) index \( k \in \{1, \ldots, m-1\} \) such that
\[
u_{j\ell}^{k(j, \ell)} = u_{j\ell}^1 \lor \cdots \lor u_{j\ell}^{m-1}.
\]
For each \( j, k \), let \( L(j, k) = \{ \ell : k(j, \ell) = k \} \) (roughly speaking, \( L(j, k) \) is the set of coordinates \( \ell \) for which the maximum of \( u_{j\ell}^1, \ldots, u_{j\ell}^{m-1} \) is attained in set \( V_k \)). Then,
\[
c(A_m) = c(A_{m-1}) + c(V_m) - \sum_{j=1}^n \sum_{\ell=1}^p \left( u_{j\ell}^{k(j, \ell)} \land u_{j\ell}^m \right) = c(A_{m-1}) + c(V_m) - \sum_{j=1}^n \sum_{k=1}^{m-1} \sum_{\ell \in L(j, k)} (u_{j\ell}^k \land u_{j\ell}^m). \tag{3.16}
\]

Consider now the quantity \( Q = \sum_{j=1}^n \sum_{k=1}^{m-1} \sum_{\ell \in L(j, k)} (u_{j\ell}^k \land u_{j\ell}^m) \). Intuitively, \( c(V_m) - Q \) in Eq. (3.16) represents the amount by which the cost of the partial solution \( A_{m-1} \) increases when the set \( V_m \) is appended to this partial solution: so, \( Q \) can be viewed as the amount of \( c(V_m) \) that is "covered" by \( V_1, \ldots, V_{m-1} \).

Clearly, there exists an index \( k^* \in \{1, \ldots, m-1\} \) such that
\[
\sum_{j=1}^n \sum_{\ell \in L(j, k^*)} (u_{j\ell}^{k^*} \land u_{j\ell}^m) \geq \frac{1}{m-1} Q
\]

(there is a set \( V_k \) that, by itself, covers at least the fraction \( \frac{1}{m-1} Q \) of the amount of \( c(V_m) \) that is covered by \( V_1, \ldots, V_{m-1} \) together).

Assume now that \( A_m \) is an optimal assignment: \( c(A_m) = c_{m}^{OPT} \). Denote by \( H_m \) the assignment produced by a sequential heuristic which optimally matches the partial assignment \( H_{m-1} \) with \( V_m \), and denote by \( H_{m,k^*} \) the assignment obtained by concatenating \( H_{m-1} \) with the assignment \( \{ (u_{j\ell}^{k^*}, u_{j\ell}^m) : j = 1, \ldots, n \} \) extracted from the optimal solution \( A_m \). Clearly, \( c(H_m) \leq c(H_{m,k^*}) \). Inequality (3.4) implies that
\[
c(H_m) \leq c(H_{m,k^*}) \leq c(H_{m-1}) + c(V_m) - \sum_{j=1}^n \sum_{\ell \in L(j, k^*)} (u_{j\ell}^{k^*} \land u_{j\ell}^m) \leq c(H_{m-1}) + c(V_m) - \frac{1}{m-1} Q.
\]

Using the definition (3.16) of \( Q \), we obtain
\[
c(H_m) \leq c(H_{m-1}) + c(V_m) - \frac{1}{m-1} \left( c(A_{m-1}) + c(V_m) - c_{m}^{OPT} \right) \tag{3.17}
\]
\[
= c(H_{m-1}) + \frac{m-2}{m-1} c(V_m) + \frac{1}{m-1} (c_{m}^{OPT} - c(A_{m-1})). \tag{3.18}
\]
Note that the inequality (3.17)-(3.18) is valid for any sequential heuristic. But we are going to apply it next to the analysis of the heaviest-first heuristic.

3.4.2. A bound for the heaviest-first heuristic

As described in Algorithm 5, the heaviest-first heuristic arises when the first assignment contains the heaviest set \( V_i \). Here, we assume with loss of generality that \( V_1 \) is the heaviest set.

We let \( \mathcal{H}^{odd}(m) = \sum_{k=1}^{m} \frac{1}{2k-1} \). Then \( \mathcal{H}^{odd}(m) = \mathcal{H}(2m-1)-\frac{1}{2}\mathcal{H}(m-1) \), where \( \mathcal{H}(m) = \sum_{k=1}^{m} \frac{1}{k} \) is the harmonic function. It is well-known that \( \ln(m+1) \leq \mathcal{H}(m) \leq 1 + \ln m \) for all \( m \geq 1 \). Thus, the function \( \mathcal{H}^{odd} \) grows like \( \frac{1}{2} \ln(m) \) and \( \mathcal{H}^{odd}(m) \geq \frac{1}{2} \ln(m) \).

**Theorem 5.** The heaviest-first heuristic \( H^{heavy} \) is a \( \frac{1}{2}(m - \mathcal{H}^{odd}(m - 1) + 1) \)-approximation algorithm for MVA-m when the cost function \( c \) is additive. Thus, \( \rho^{heavy}(m) \leq \frac{1}{2}(m - \mathcal{H}^{odd}(m - 1) + 1) \leq \frac{1}{2}(m + 1) - \frac{1}{2} \ln(m - 1) \).

**Proof.** Let \( H_m \) be the solution found by the heaviest-first heuristic. The proof proceeds by induction, starting with \( m = 2 \) and \( \rho^{heavy}(2) = 1 \).

Consider first the case where \( c(V_1) \leq \frac{m-1}{2m-3} \cdot c^{OPT} \). By induction,

\[
c(H_{m-1}) \leq \rho^{heavy}(m - 1) \cdot c^{OPT}_{m-1},
\]

where \( c^{OPT}_{m-1} \) is the cost of the optimal assignment for \( V_1 \times \ldots \times V_{m-1} \). Let \( A_m \) again be an optimal assignment for \( V_1 \times \ldots \times V_m \). Clearly, \( c^{OPT}_{m-1} \leq c(A_{m-1}) \). Using this in (3.17) together with \( c(V_m) \leq c(V_1) \leq \frac{m-1}{2m-3} c^{OPT} \) yields

\[
c(H_m) \leq \rho^{heavy}(m - 1) c(A_{m-1}) + \left( \frac{m-2}{2m-3} \right) \cdot c^{OPT}_{m-1} + \frac{1}{m-1} \left( c^{OPT}_{m-1} - c(A_{m-1}) \right).
\]

The alternative case is when \( c(V_1) \geq \frac{m-1}{2m-3} \cdot c^{OPT} \). Repeat the analysis leading to Eq. (3.15) in the second part of the proof of Theorem 4, but this time with \( V_1 \) replacing \( V_{m-1} \). From there,

\[
c(H_m) \leq c(H_{m-1}) + c(M_{1,m}) - c(V_1)
\]

\[
\leq \rho^{heavy}(m - 1) c^{OPT}_{m-1} + c_{m}^{OPT} - \frac{m-1}{2m-3} c^{OPT}_{m}
\]

\[
\leq \left( \rho^{heavy}(m - 1) + \frac{m-2}{2m-3} \right) c^{OPT}_{m}.
\]

 Altogether, we obtain the recurrence equation:

\[
\rho^{heavy}(m) = \rho^{heavy}(m - 1) + \frac{m-2}{2m-3}.
\]

To analyze this relation, let \( r_m = m - 2 \rho^{heavy}(m) \). Then, \( r_m - r_{m-1} = \frac{1}{2k-1} \), so that \( r_m = r_2 + \sum_{k=2}^{m-1} \frac{1}{2k-1} \). Since \( r_2 = 0 \), \( r_m = \mathcal{H}^{odd}(m - 1) - 1 \). \( \Box \)
The tightness of the bound established in Theorem 5 is discussed in Section 3.5.

3.5. Bad instances for additive cost functions

In this section we complement the previous results by showing that hub and sequential algorithms can perform rather poorly even when the cost function is additive. (Recall that for monotone submodular nonadditive functions, the bounds in Theorem 3 and Theorem 4 were already shown to be tight for all \( m \geq 2 \), even for the multi-hub and for the heaviest-first heuristic.)

Let us first consider the case \( m = 3 \). For MVA-3 with additive costs, Dokka et al. [5] established the validity and the tightness of the bounds established in Theorem 4 and Theorem 5, respectively. To see the former, observe that tightness of the bound \( \rho^{seq}(3) = \frac{3}{2} \) follows from the instance depicted in Figure 1.1: indeed, for this instance, \( c^{OPT} = 2 \), whereas the sequential heuristic might find a solution with value 3.

To see that \( \rho^{heavy}(3) = \frac{4}{3} \), consider the instance with \( p = 3 \), \( V_1 = \{ e_1, e_2, 0 \} \), \( V_2 = \{ e_3, e_2, 0 \} \), \( V_3 = \{ e_1, 0, e_3 \} \). Its optimal value is \( c^{OPT} = 3 \), whereas \( H^{heavy} \) might produce first \( H_2 = \{ (e_1, e_3), (e_2, e_2), (0, 0) \} \), then \( H_3 = \{ (e_1, e_3, e_1), (e_2, e_2, 0), (0, 0, e_3) \} \), with \( c(H_3) = 4 \).

An obvious improvement to heuristics \( H^{seq} \) and \( H^{heavy} \) would be to run \( H^{seq} \) for all possible permutations of the sets \( V_1, \ldots, V_m \) in the first step, then to retain the best of the \( m! \) feasible solutions found (see Bandelt et al. [1], Crama and Spieksma [4] for related “multiple-pass” heuristics). Interestingly, when \( m = 3 \), it follows again from the previous example that this multiple-pass heuristic (which involves solving six bipartite matching problems) has the same worst-case ratio as \( H^{heavy} \) (which only solves two matching problems). This observation also entails that the ratio \( \rho(3) = \frac{4}{3} \) is tight for the iterative matching algorithm of Reda et al. [11].

Let us now turn to the general case \( m \geq 3 \) for additive cost functions. The ratio \( \rho^{hub}(m) = \frac{m}{2} \) is tight in this case for the heaviest-hub heuristic, as illustrated by the following example: Let \( p = 2 \) and \( n = m \), let \( V_1 \) contain \( e_1 \) and let \( V_2, \ldots, V_m \) contain \( e_2 \); all other vectors are \( 0 \). Then, \( c^{OPT} = 2 \) but the heaviest-hub heuristic may yield \( c(H^{hub}(V_i)) = m \). For multi-hub, on the other hand, Dokka et al. [5] give an example showing that the performance ratio of the heuristic may be as bad as \( \frac{m}{2} \), whereas Theorem 3 only proves the upper bound \( \frac{m}{2} \). We do not know the exact approximation ratio of multi-hub for additive cost functions.

Dokka et al. [5] observed that the worst-case approximation ratio of the sequential heuristic can grow as fast as \( \Omega(\sqrt{m}) \) for certain instances with additive cost functions. We now strengthen this result by establishing a lower bound of the same order for the heaviest-first heuristic.

**Theorem 6.** There exists an infinite sequence of values of \( m \) such that the heaviest-first heuristic produces a feasible assignment with cost larger than \( \frac{\sqrt{m}}{2} c^{OPT} \) on certain instances of WWI-m.

**Proof.** Fix an arbitrary positive integer \( r \). We are going to describe an instance of WWI-m with \( m = r^2 + 1 \) and \( n = p = 2r \). In order to simplify the description of the instance, we label the input sets from \( V_0 \) to \( V_{2r} \). We write \( v_{ij} \) to denote the \( j^{th} \) vector of set \( V_i \), \( i = 0, \ldots, r^2 \), \( j = 1, \ldots, 2r \). The construction of the sets \( V_0, V_1, \ldots, V_{2r} \) is as follows. (An instance with \( r = 3 \) is displayed in Figure 3, and the corresponding heuristic and optimal solutions are illustrated in Figure 4.)
In $V_0$, $v_{0j} = e_j$ for $j = 1, \ldots, r$, and $v_{0j} = \mathbf{0}$ for $j = r + 1, \ldots, 2r$.

For $i > 0$, write $i = (k - 1)r + \ell$ with $k, \ell \in \{1, \ldots, r\}$. Then, in $V_i$,

- for $j = 1, \ldots, r$, $v_{ij} = e_j$ if $j \neq \ell$ and $v_{i\ell} = \mathbf{0}$;
- for $j = r + 1, \ldots, 2r$, $v_{ij} = \mathbf{0}$ if $j \neq r + k$ and $v_{i,r+k} = e_{r+k}$.

For this instance, the optimal cost equals $2r$: for $j = 1, \ldots, 2r$, the $j$th tuple of the optimal assignment simply collects all vectors $e_j$ (note that there is at most one such $e_j$ in each set $V_i$).

However, the heaviest-first heuristic may find a solution with cost $r^2 + r$ as follows: First, note that $c(V_i) = r$ for all $i$, so that $H_{\text{heavy}}$ may consider the sets $V_0, V_1, \ldots, V_{2r}$ in that order. When matching $V_1$ to $V_0$, $H_{\text{heavy}}$ may assign the $(r + 1)^{st}$ vector of $V_1$ to the first vector of $V_0$. In the next $r - 1$ assignment stages, it assigns the $(r + 1)^{st}$ vector of $V_i$ ($i = 2, \ldots, r$) to the tuple containing the $i^{th}$ vector of $V_0$. Then, in the next $r$ assignments, $H_{\text{heavy}}$ assigns the $(r + 2)^{nd}$ vector of $V_i$ ($i = r + 1, \ldots, 2r$) to the tuple $i = 1, \ldots, r$ containing the $i^{th}$ vector of $V_0$. Proceeding in this way yields a solution with cost $r^2 + r$.

For $r = 3$ and $m = 10$, the optimal assignment is shown in Figure 3:

![Figure 3: A bad instance for the heaviest-first heuristic with $r = 3, m = 10$](image-url)

### 4. WWI-3 is hard to approximate

As mentioned earlier, Reda et al. [11] have observed that WWI-$m$ is NP-hard for $m \geq 3$. An explicit proof is found in Dokka et al. [5]. Our objective is now to strengthen this result by showing that WWI-3 does not admit a polynomial-time approximation scheme, unless P=NP.
We shall describe a reduction from 3-bounded maximum 3-dimensional matching (MAX-3DM-3) to WWI-3. An instance of MAX-3DM-3 consists of three pairwise disjoint sets $X, Y, Z$ such that $|X| = |Y| = |Z| = q$, and of a set of triples $S \subseteq X \times Y \times Z$ such that every element of $X \cup Y \cup Z$ appears in at most three triples of $S$; let $|S| = s$. A matching in $S$ is a subset $S' \subseteq S$ such that no element of $X \cup Y \cup Z$ appears in two triples of $S'$. The goal of the MAX-3DM-3 problem is to find a matching of maximum cardinality in $S$.

Kann [8] showed that MAX-3DM-3 is APX-hard. An instance of MAX-3DM-3 is called a perfect instance if its optimal solution consists of $q$ triples that cover all elements of $X \cup Y \cup Z$ (that is, if $S$ contains a feasible assignment). Petrank [10] proved that perfect instances of MAX-3DM-3 are hard to approximate, and that the existence of a polynomial-time approximation scheme for perfect instances would imply $P=NP$.

Now, consider an arbitrary perfect instance $I'$ of MAX-3DM-3. We build a corresponding instance $I$ of WWI-3 by using the gadget depicted in Figure 5, as explained next.

The instance $I$ consists of three sets $V_X, V_Y, V_Z$, each of cardinality $q + 3s$. Each element $e$ of each $V_k, k \in \{X, Y, Z\}$, is a 0-1 vector of length $6q+4s$ containing exactly two nonzero elements. So, we can view each $e$ as an edge in an undirected graph $G = (U, A)$ where $U$ is a vertex set with cardinality $6q+4s$ and $A$ can be identified with $V_X \cup V_Y \cup V_Z$. The elements of $U$ are

- $x_1, x_2$ for each $x \in X$
- $y_1, y_2$ for each $y \in Y$
- $z_1, z_2$ for each $z \in Z$

Figure 4: Cost of optimal solution is 6; cost returned by $H_{\text{heavy}}$ is 12
• $x_t, y_t, z_t$ and $u_t$ for each triple $t \in S$

and the edges in $A = V_X \cup V_Y \cup V_Z$ are

• $(x_1, x_2) \in V_X$ for each $x \in X$ (element edges)
• $(y_1, y_2) \in V_Y$ for each $y \in Y$ (element edges)
• $(z_1, z_2) \in V_Z$ for each $z \in Z$ (element edges)
• $(x_1, x_t) \in V_Y, (x_2, x_t) \in V_Z, (y_1, y_t) \in V_X, (y_2, y_t) \in V_Z, (z_1, z_t) \in V_X, (z_2, z_t) \in V_Y$, for each $t \in S$ (gadget edges)
• $(x_t, u_t) \in V_X, (y_t, u_t) \in V_Y, (z_t, u_t) \in V_Z$ for each $t \in S$ (gadget edges).

We say that an element of $V_X$ ($V_Y, V_Z$) is an $X$-edge ($Y$-edge, $Z$-edge). The subgraph induced by all gadget edges associated with a same triple $t$ is called the gadget associated with $t$ and is denoted by $g(t)$. Note that $g(t)$ contains three element edges.

Observe that a feasible triple for $WIII-3$ consists of an $X$-edge, a $Y$-edge and a $Z$-edge. A feasible triple of edges $T \subseteq A$ defines (and can be identified with) a subgraph $(U_T, T)$ of $G$, where $U_T$ is the subset of vertices covered by $T$. The cost of $T$ is $|U_T|$. Note that a feasible triple is either

• a triangle $K_3$ with cost 3, or
• a claw $K_{1,3}$, or a path $P_4$, with cost 4, or
• disconnected with cost either 5 or 6.

We say for short that $T$ is connected if $(U_T, T)$ is connected.

A feasible assignment for $I$ is a collection of $q + 3s$ feasible triples covering all edges of $G$. We now collect some properties of feasible assignments for further reference.
Lemma 14. Let \( M \subseteq V_X \times V_Y \times V_Z \) be a feasible assignment for \( I \), with \( |M| = q + 3s \).

(1) \( M \) contains at most \( 3q \) triangles.

(2) The cost of \( M \) (and hence, the optimal value of \( I \)) is at least \( q + 12s \).

(3) If the cost of \( M \) is \( q + 12s \), then \( M \) contains \( 3q \) triangles, \( 3s - 2q \) additional connected triples, and no disconnected triples.

(4) If the cost of \( M \) is equal to \( q + 12s + r \) \((r \geq 0)\), then \( M \) contains at least \( 3q - r \) triangles and at most \( r \) disconnected triples.

Proof. (1) We say \( M \) covers \( A' \), with \( A' \subseteq A \), if all edges in \( A' \) are contained in \( M \). Observe that \( M \) contains the same number of edges as \( A \) (namely, \( 3q + 9s \) edges), and hence, since \( M \) covers \( A \), each edge of \( A \) must be covered exactly once. In particular, each element edge can be covered by at most one triangle, which implies that there are at most \( 3q \) triangles in \( M \).

(2) The cost of \( M \) is equal to \( 3c_3 + 4c_4 + 5c_5 + 6c_6 \), where \( c_k \) is the number of triples with cost equal to \( k \), and \( c_3 + c_4 + c_5 + c_6 = |M| \). There holds:

\[
c(M) = 3c_3 + 4c_4 + 5c_5 + 6c_6 \quad (4.19)
\]

\[
\geq 3c_3 + 4(|M| - c_3 - c_5 - c_6) + 5(c_5 + c_6) \quad (4.20)
\]

\[
= -c_3 + (c_5 + c_6) + 4|M|. \quad (4.21)
\]

Since \( c_3 \leq 3q \) and \( c_5 + c_6 \geq 0 \), Eq. (4.21) implies that the cost of \( M \) is at least \(-3q + 4(q + 3s) = q + 12s \).

(3) The previous reasoning shows that the cost of \( |M| \) can be equal to \( q + 12s \) only if \( c_3 = 3q \) and \( c_5 + c_6 = 0 \).

(4) Intuitively, every missing triangle and every disconnected triple increases the cost of \( M \) by at least one unit with respect to the lower bound \( q + 12s \), as expressed by the inequality (4.21). More formally, if \( c_3 < 3q - r \), then Eq. (4.21) leads to

\[
3c_3 + 4c_4 + 5c_5 + 6c_6 > -(3q - r) + 4|M| = q + 12s + r. \quad (4.22)
\]

Similarly, if \( c_5 + c_6 > r \), then Eq. (4.21) together with \( c_3 \leq 3q \) imply

\[
3c_3 + 4c_4 + 5c_5 + 6c_6 > -3q + r + 4|M| = q + 12s + r. \quad (4.23)
\]

We are now ready to establish the relation between the solutions of \( I \) and \( I' \)

Lemma 15. If \( I' \) is a perfect instance of MAX-3DM-3, then the optimal value of \( I \) is \( q + 12s \).

Proof. If \( t \in S \) is in the perfect matching, then use three triangles and the claw centered at \( u_t \) in the associated gadget \( g(t) \). Otherwise, use three claws centered at \( x_t, y_t \) and \( z_t \), respectively. Clearly, in the constructed solution for WWI-3 there are only triangles and claws with exactly \( 3q \) triangles. Hence, by Lemma 14 it follows that the cost of the solution is \( q + 3s \).

The converse statement will follow from Lemma 16 hereunder, with \( \delta \geq 0 \).
Lemma 16. Let $\delta \geq 0$ be a real number. If instance $I$ has a feasible solution with cost at most $q + 12s + \delta q$, then instance $I'$ possesses a matching with size at least $(1 - 6\delta)q$.

Proof. Consider a feasible solution $M$ for instance $I$ with cost at most $q + 12s + \delta q$. We call a gadget damaged (by $M$) if:

(Type (g)) at least one of its gadget edges is in a disconnected triple of $M$, or

(Type (e)) one of its element edges is not included in a triangle of $M$.

Equivalently, a gadget is undamaged if all its gadget edges are in connected triples of $M$ and if all its element edges are in triangles of $M$.

We call an element edge damaged (by $M$) if it is not included in a triangle of $M$, or it is in a triangle contained in a damaged gadget. Equivalently, an element edge is undamaged if it is in a triangle contained in an undamaged gadget.

It follows from Lemma 14 that $M$ contains at least $3q - \delta q$ triangles. Thus, at most $\delta q$ element edges are not included in triangles.

Note that if an edge is damaged, then it is contained in a damaged gadget. Since $I'$ is an instance of MAX-3DM-3, each element edge occurs in at most three gadgets. In particular, each damaged element edge can damage at most three gadgets, so that there are at most $3\delta q$ damaged gadgets of type (e).

Furthermore, Lemma 14 also implies that at most $\delta q$ triples can be disconnected; these triples contain at most $3\delta q$ gadget edges, which can damage at most $3\delta q$ gadgets (damaged gadgets of type (g)).

Since each damaged gadget may yield at most three damaged element edges of type (ii), we find that, altogether there are at most $18\delta q$ damaged element edges, which leaves at least $3(1 - 6\delta)q$ undamaged element edges.

counting of $\delta q$ potential damaged edges of type (i).)

Now, the main element of the proof of the lemma is the following claim:

Claim 17. Every undamaged element edge, say $(x_1, x_2)$, is in a triangle $(x_1, x_2, x_3)$ from some undamaged gadget $g(t)$. We claim that the other two element edges in $g(t)$ are also included in triangles from $g(t)$.

(Proof of claim.) To see this, consider one of the other element edges in $g(t)$, say $(y_1, y_2)$. Since $g(t)$ is undamaged, $(y_1, y_2)$ must be covered by a triangle contained in a gadget $g(t')$. Assume by contradiction that $t \neq t'$ (otherwise, we are done).

Again because $g(t)$ is undamaged, the $X$-edge $(y_1, y_1)$ is in a connected triple $T$, which must necessarily contain the $Y$-edge $(y, u)$ (indeed, at vertex $y_1$, $(y_1, y_1)$ is only incident to $X$-edges and to the $Y$-edge $(y_1, y_2)$ which is already covered by a triangle in $g(t')$; so, $T$ must contain either the $Y$-edge $(y_1, u)$ or the $Z$-edge $(y_2, y)$; but the latter case implies the former one).

The previous reasoning applies similarly to $(y_2, y)$, so that the claw $\{(y_1, y_1), (y_2, y_1), (y_1, u_1)\}$ must be in $M$.

This implies, in turn, that $(x_1, u_1)$ and $(z_1, u_1)$ must be in the same triple, which can only contain $(z_2, z_1)$ as $Y$-edge. Thus, $(z_1, z_2)$ must be in a triple together with $(z_1, z_2)$, contradicting the hypothesis that $(z_1, z_2)$ is undamaged. (End of claim.)

Hence the $3(1 - 6\delta)q$ undamaged element edges can be divided into groups of three that correspond to $(1 - 6\delta)q$ undamaged gadgets. Then the corresponding $(1 - 6\delta)q$ triples in instance $I'$ form a matching. ■

We are now ready for the main result of this section.
Theorem 7. WWI-3 is APX-hard even when all vectors in $V_X \cup V_Y \cup V_Z$ are 0-1 vectors with exactly two nonzero entries per vector.

Proof. When we apply the reduction to a perfect instance $I'$ of MAX-3DM-3, Lemma 15 yields $c^{OPT}(I) = q + 12s$ for the resulting instance $I$ of WWI-3. A $(1 + \epsilon)$-approximation algorithm for WWI-3 would imply that we can compute, in polynomial time, a solution of $I$ with objective value at most equal to

$$(1 + \epsilon) c^{OPT}(I) \leq q + 12s + 37\epsilon q$$

(here we have used $s \leq 3q$). Then Lemma 16 (with $\delta = 37\epsilon$) implies the existence of a matching of size at least $(1 - 222\epsilon)$ for instance $I'$, and this matching can be found in polynomial time. Hence, a PTAS for WWI-3 would imply a PTAS for any perfect instance of 3-bounded MAX-3DM. ■

5. Binary inputs and fixed $p$

In this section we consider again the binary MVA problem, that is, the special case of MVA where all vectors in $V_1 \cup \ldots \cup V_m$ are binary. We want to argue that the binary MVA problem can be solved in polynomial time when $p$ is fixed.

For an instance of the binary MVA problem, we let, as in Theorem 6, $v_{ij}$ denote the $j^{th}$ vector in set $V_i$, $j = 1, \ldots, n$, $i = 1, \ldots, m$. Let $b_1, \ldots, b_{2p}$ be all distinct 0-1 vectors of length $p$, arbitrarily ordered, and consider a feasible $m$-tuple $(u^1, \ldots, u^m)$. We say that $(u^1, \ldots, u^m)$ is of type $t$ if $u^1 \lor \ldots \lor u^m = b_t$.

We construct a mixed integer formulation of MVA featuring variables $x_t$:

$$x_t = \text{number of } m\text{-tuples of type } t \text{ in the assignment, } t = 1, \ldots, 2^p.$$ 

We also need assignment variables: for each $i = 1, \ldots, m; j = 1, \ldots, n; t = 1, \ldots, 2^p$,

$$z_{ij}^t = 1 \text{ if } v_{ij} \text{ is assigned to an } m\text{-tuple of type } t.$$ 

The formulation is now:

$$\min \sum_{t=1}^{2^p} c(b_t)x_t \quad (5.24)$$

$$\sum_{j: b_t \geq v_{ij}} z_{ij}^t = x_t \quad \text{for each } t = 1, \ldots, 2^p, i = 1, \ldots, m, \quad (5.25)$$

$$\sum_{t: b_t \geq v_{ij}} z_{ij}^t = 1 \quad \text{for each } j = 1, \ldots, n, i = 1, \ldots, m, \quad (5.26)$$

$$x_t \text{ integer for each } t = 1, \ldots, 2^p, \quad (5.27)$$

$$z_{ij}^t \geq 0 \text{ for each } j = 1, \ldots, n, t = 1, \ldots, 2^p, i = 1, \ldots, m. \quad (5.28)$$

The objective function (5.24) minimizes the total cost. Constraints (5.25)-(5.26) are the familiar transportation constraints. Notice further that integrality of $x_t$ implies integrality of $z_{ij}^t$. 

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Lemma 18. Formulation (5.24)-(5.28) is a correct formulation of the binary MVA problem.

Proof. Consider a feasible solution of the binary MVA problem. This solution prescribes, for each binary vector \( v_{ij} \) in each set \( V_i \), whether this vector should be assigned to an \( m \)-tuple of type \( t \). This determines the \( x_t \) and \( z_{jt} \) values, which clearly satisfy constraints (5.25)-(5.28).

Conversely, consider \( x_t, z_{jt} \) values that satisfy (5.25)-(5.28). One can construct a feasible solution of MVA-\( m \) as follows: (1) Create a set \( X \) containing a copy of vector \( b_t \) for each \( x_t > 0 \). (2) For each \( i = 1, \ldots, m \), construct a bipartite graph \( G = (V_i \cup X, E) \) where vector \( v_{ij} \) of \( V_i \) is connected with vector \( b_t \) of \( X \) if \( v_{ij} \leq b_t \). The values \( x_t, z_{jt} \) define a feasible solution of the transportation problem with supply equal to 1 for each vertex in \( V_i \) and demand equal to \( x_t \) for vertex \( t \) in \( X \). (3) Construct \( m \)-tuples of vectors by assigning \( m \) vectors – one from each \( V_i \) – to the same \( m \)-tuple if they all are matched to same vector in \( X \) in the solution of the transportation problem (there may be several ways of performing this step; however, any way suffices). This yields a feasible solution of the MVA problem with value at most equal to \( \sum_{t=1}^{2^p} c(b_t)x_t \). Hence, the optimal value of (5.24)-(5.28) is equal to the optimal value of the MVA problem.

Theorem 8. Binary MVA can be solved in polynomial time for each fixed \( p \).

Proof. Lemma 18 shows that formulation (5.24)-(5.28) is correct. This formulation involves \( 2^p \) integer variables \( x_t \), \( O(mn2^p) \) continuous variables \( z_{jt} \), and \( O(m2^p + mn) \) constraints. When we fix \( p \), this results in a fixed number of integer variables \( x_t \), each of which takes at most \( n + 1 \) distinct values. Therefore, in order to find an optimal solution it is enough to check the feasibility of (5.25)-(5.28) for \( O(n2^p) \) assignments of values to the \( x_t \) variables, and to choose the solution with the minimum cost.

6. Conclusions

In this paper, we have considered the multi-dimensional vector assignment problem MVA-\( m \) and we have analyzed the performance of several polynomial-time heuristics for this problem in terms of their worst-case approximation ratio. We have also proved that the problem is APX-hard, even when \( m = 3 \). Among the main questions that remain open at this stage, let us mention the following ones:

1. What is the exact approximation ratio of the multi-hub heuristic in case of additive costs? We know that it lies between \( m/4 \) and \( m/2 \).
2. What is the exact approximation ratio of the heaviest-first sequential heuristic in case of additive costs? We know that it lies between \( \Omega(\sqrt{m}) \) and \( O(m - \ln m) \).
3. Does there exist a polynomial-time algorithm with constant (i.e., independent of \( m \)) approximation ratio for MVA-\( m \)?

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