

Limited-Magnitude Error-Correcting Gray Codes for Rank Modulation

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Abstract—We construct Gray codes over permutations for the rank-modulation scheme, which are also capable of correcting errors under the infinity-metric. These errors model limited-magnitude or spike errors, for which only single-error-detecting Gray codes are currently known. Surprisingly, the error-correcting codes we construct achieve a better asymptotic rate than that of presently known constructions not having the Gray property, and exceed the Gilbert-Varshamov bound. Additionally, we present efficient ranking and unranking procedures, as well as a decoding procedure that runs in linear time. Finally, we also apply our methods to solve an outstanding issue with error-detecting rank-modulation Gray codes (snake-in-the-box codes) under a different metric, the Kendall τ -metric, in the group of permutations over an even number of elements S_{2n} , where we provide asymptotically optimal codes.

Index Terms—Gray codes, error-correcting codes, permutations, spread- d circuit codes, rank modulation

I. INTRODUCTION

RANK modulation is a method for storing information in non-volatile memories [21], which has been researched in recent years. It calls for storing information in relative values stored in a group of cells rather than the absolute values of single cells. More precisely, it stores information in the permutation suggested by sorting a group of cells by their relative values, e.g., charge levels in flash memory cells. It allows for increased robustness against certain noise mechanisms (e.g., charge leakage in flash memory cells), as well as alleviating some inherent challenges in flash memories (e.g., programming/erasure-asymmetry and programming-overshoot). Permutation codes have also previously seen usages in source-encoding [3]–[5], [31] and signal detection [7], as well as other fields [6], [9], [10], and more recently been used in power-line communications [34].

Several error models have been studied for rank modulation, including the Kendall τ -metric [2], [22], [25], [39], the ℓ_∞ -metric [24], [30], [32], [33] and other examples [11], [16]. In this paper we focus on the ℓ_∞ -metric, which models limited-magnitude or spike noise, i.e., we assume that the rank of any given cell—its position when sorting the group of cells—could not have changed by more than a given amount. [24], [32] have presented constructions for error-correcting codes under this metric, as well as explored some non-constructive lower- and upper-bounds on the parameters of existing codes. [33] has

since employed methods of relabeling to optimize the minimal distance of known constructions.

Gray codes were first discussed over the space of binary vectors, where each pair of consecutive vectors differed by a single bit-flip [17]; the concept has since been generalized to include codes over arbitrary alphabets, requiring only that codewords could be ordered in a sequence, where each codeword is derived from the previous by one of a predefined set of functions. Other suggested usages of Gray codes, surveyed in [28], include permanent-computation [26], circuit-testing [27], image-processing [1], hashing [15], coding [14], [21], [29] and data storing/extraction [8]. In particular, in the context of rank modulation, the use of Gray codes has been shown to reduce write-time by eliminating the risk of programming-overshoot and allow integration with other multilevel-cells coding schemes [12], [13], [21].

Gray codes with error-correction capabilities have sometimes been referred to as spread- d circuit codes (see [19] and references therein). Specifically, in the context of rank modulation, such codes were so far only studied for the case of single-error-detection, where they were dubbed snake-in-the-box codes (or, more appropriately, coil-in-the-box codes, when they are cyclic). [36] studied these codes under both the Kendall τ -metric and the ℓ_∞ -metric, and more recent papers [18], [20], [38] have categorized and constructed optimally sized coil-in-the-box codes under the former metric for odd orders, although the case of even orders proved more challenging (see [37] in addition to the aforementioned papers).

In this work we focus on the ℓ_∞ -metric and present a construction of error-correcting Gray codes capable of correcting an arbitrary number of limited-magnitude errors. The allowed transitions between codewords are the “push-to-the-top” operations, used in most previous works [12], [13], [18], [20], [21], [36], [38]. The resulting codes will be shown to be larger than known constructions in the case of fixed minimal distance, as well as achieve better asymptotic rates than known codes in the case of $d = \Theta(n)$.

We will also briefly examine error-detecting codes under the Kendall τ -metric for even orders, since methods developed for the application of our main construction can readily be adapted to that purpose. We provide an asymptotically optimal construction which nearly completes the categorization of available codes for that scheme.

The paper is organized as follows. In Section II we present notations and definitions. In Section III we study a new kind of auxiliary codes which are required for our construction, before presenting it in Section IV and discussing its performance in comparison with known constructions and bounds. We devise

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TABLE I
CODE NOTATIONS FOR $C \subseteq S_n$.

Notation	Definition
$G_{i\uparrow}(n, M)$	$C = (c_r)_{r=1}^M \subseteq S_n$ such that for all r : $c_{(r \bmod M)+1} = t_{i\uparrow r}(c_r)$.
$G_{\uparrow}(n, M)$	C is a $G_{1\uparrow}(n, M)$.
$G_{\downarrow}(n, M)$	$C = (c_r)_{r=1}^M \subseteq S_n$ such that for all r : $c_{(r \bmod M)+1} = t_{\downarrow jr}(c_r)$.
(n, M, d) -LMRM	$C \subseteq S_n$, $ C = M$ and for all $c_1 \neq c_2 \in C$: $d_{\infty}(c_1, c_2) \geq d$.
$G_{i\uparrow}(n, M, d)$	C is a $G_{i\uparrow}(n, M)$ and an (n, M, d) -LMRM.
$G_{\uparrow}(n, M, d)$	C is a $G_{1\uparrow}(n, M, d)$.
$G_{\uparrow}^{\text{aux}}(k, M)$	C is a $G_{\uparrow}(k, M)$, and for all $q \in [k-1]$: $\sigma \in C \implies (q, k)\sigma \notin C$. (See Section III.)

codes, but the codes presented in this paper are nevertheless also, in particular, spread- d circuit-codes.

We shall focus on the ℓ_{∞} -metric defined on S_n by

$$d_{\infty}(\sigma, \tau) = \max_{j \in [n]} |\sigma(j) - \tau(j)|.$$

That is, it is the metric induced on S_n by the embedding into \mathbb{Z}^n (and, indeed, \mathbb{R}^n) implied by the vector notation, and the ℓ_{∞} -metric in these spaces. [32] studied error-correcting codes in S_n with d_{∞} , which it dubbed *limited-magnitude rank-modulation* codes, and denoted a code C with minimal distance d as an $(n, |C|, d)$ -LMRM code. In our case, if a $G_{i\uparrow}(n, M)$ is also an (n, M, d) -LMRM code, we shall denote it a $G_{i\uparrow}(n, M, d)$ (likewise for G_{\uparrow} and G_{\downarrow}).

Finally, we organize our notation of codes in Table I.

III. AUXILIARY CONSTRUCTION

Before we present the main construction of our paper, we first describe in this section a construction for auxiliary codes which will be a component of the main construction.

We define auxiliary codes in S_k in the following way: we say that C is $G_{\uparrow}^{\text{aux}}(k, M)$ if it is $G_{\uparrow}(k, M)$ and for all $q \in [k-1]$ it holds that

$$\sigma \in C \implies (q, k) \circ \sigma \notin C.$$

In our main construction, we will use a $G_{\uparrow}^{\text{aux}}(k, M)$ code $(c_r)_{r=1}^M$, which we also require to satisfy $c_1 = \text{Id}$ and $c_2 = t_{\uparrow k} \text{Id}$. We hence study the existence of such codes.

Firstly, note that the only existing $G_{\uparrow}^{\text{aux}}(2, M)$ codes are the singletons $\{\text{Id}\}, \{(1, 2)\}$. However, for $k \geq 3$ there do exist $G_{\uparrow}^{\text{aux}}(k, M)$ codes with $M \geq 3$, as one such example is

$$\left\{ \text{Id}, t_{\uparrow 3} \text{Id}, t_{\uparrow 3}^2 \text{Id} \right\}.$$

We also note the following:

Lemma 3 If $C \subseteq S_k$ is $G_{\uparrow}^{\text{aux}}(k, M)$, then $M \leq \frac{|S_k|}{2}$.

Proof: Take $q \in [k-1]$, and observe that $\sigma \mapsto (q, k)\sigma$ is an S_k -automorphism, under which C and its image are disjoint. Hence $2M \leq |S_k|$. ■

This motivates us to examine another family of codes, namely, parity-preserving codes, due to the following observations.

Lemma 4 If $C \subseteq S_k$ is parity-preserving then $|C| \leq \frac{|S_k|}{2}$.

Proof: Either $C \subseteq A_k$ or $C \subseteq S_k \setminus A_k$. ■

Lemma 5 If $C \subseteq S_k$ is a parity-preserving $G_{\uparrow}(k, M)$, then C is $G_{\uparrow}^{\text{aux}}(k, M)$.

Proof: Observe that $\text{sign } \sigma \neq \text{sign}(q, k)\sigma$ for all $q \in [k-1]$, thus both cannot belong to C . ■

Parity-preserving $G_{\uparrow}^{\text{aux}}(2m+1, M)$ codes are known to exist, achieving the aforementioned bound.

Lemma 6 [18] For all $m \neq 2$, there exist parity-preserving $G_{\uparrow}(2m+1, \frac{(2m+1)!}{2})$ codes. The largest parity-preserving $G_{\uparrow}(5, M)$ codes have $M = 57$.

Although not declared, it is shown in [18] that such codes can be assumed to have $t_{\uparrow 2m+1}$ as the first transition in their generating transition sequence, and furthermore, that they also employ at least one $t_{\uparrow 2m-1}$ transition.

In comparison, as noted in [36], a parity-preserving $G_{\uparrow}(2m, M)$ must satisfy $M \leq \frac{|S_{2m}|}{2m}$, as it must never employ a $t_{\uparrow 2m}$ transition. We therefore examine more general $G_{\uparrow}^{\text{aux}}$ codes, which are not parity-preserving. We begin by noting the following lemma.

Lemma 7 [21, Thm. 4,7] For all $n \geq 1$ there exist $G_{\uparrow}(n, n!)$ codes, that is, complete and cyclic ‘‘push-to-the-top’’ Gray codes over the symmetric group S_n .

Relying on these codes, we construct auxiliary codes in the following theorem. Similarly constructed codes already appeared in a different context as components in a construction in [21].

Theorem 8 For all $m \geq 2$ there exists a $G_{\uparrow}^{\text{aux}}(2m, \frac{|S_{2m}|}{2m-1})$, starting at Id and with a generating sequence starting with $t_{\uparrow 2m}$.

Proof: Take a $G_{\uparrow}(2m-2, (2m-2)!)$ code C' , provided by Lemma 7. We follow the concept of [21, Thm. 7] in extending C' to S_{2m} . Let us define

$$\sigma_0 = t_{\uparrow 2m} \text{Id} = [2m, 1, \dots, 2m-1].$$

If we take $t_{\uparrow i_1}, t_{\uparrow i_2}, \dots, t_{\uparrow i_{(2m-2)}}$ to be the transition sequence generating C' , then the transition sequence

$$t_{\downarrow 2m+1-i_1}, t_{\downarrow 2m+1-i_2}, \dots, t_{\downarrow 2m+1-i_{(2m-2)}}$$

of ‘‘push-to-the-bottom’’ operations, applied in succession to σ_0 , generates $C'' \subseteq S_{2m}$, a $G_{\downarrow}(2m, (2m-2)!)$, all of whose elements’ vector notations begin with $[2m, 1]$.

We now note that $t_{\downarrow 2m+1-j} = t_{\uparrow 2m}^{2m-1} t_{\uparrow j}$. Thus, by replacing each $t_{\downarrow 2m+1-j}$ with $t_{\uparrow j}$ followed by a sequence of $2m-1$ occurrences of $t_{\uparrow 2m}$, we get $C \subseteq S_{2m}$, a

$G_{\uparrow}(2m, (2m-2)!2m)$, where every block of $2m$ elements is comprised of cyclic shifts of some $\sigma \in C''$.

The code C is known to be a Gray code [21, Thm. 7]. Moreover, if $\sigma \in C$ satisfies $\tau = (q, 2m)\sigma \in C$, note that both have a vector notation with 1 immediately (cyclically) following $2m$, but since $\tau = (q, 2m)\sigma$ its vector notation has 1 following q . It follows (by abuse of notation) that $q = 2m$.

Finally, note that C is generated by a transition sequence ending with $2m-1$ instances of $t_{\uparrow 2m}$, so it includes Id followed by a $t_{\uparrow 2m}$ transition. A cyclic shift of C therefore satisfies the theorem. ■

Example 9 To construct a $G_{\uparrow}^{\text{aux}}(4, 8)$ we utilize the complete $G_{\uparrow}(2, 2)$ code $\{\text{Id}, t_{\uparrow 2} \text{Id}\}$, generated by $t_{\uparrow 2}, t_{\uparrow 2}$, to arrive by the $G_{\downarrow}(4, 2)$ code

$$C'' = \{[4, 1, 2, 3], [4, 1, 3, 2]\},$$

which is generated by $t_{\downarrow 3}, t_{\downarrow 3}$. We recall that $t_{\downarrow 3} = t_{\uparrow 4}^3 \circ t_{\uparrow 3}$, allowing us to expand C'' in the following manner:

$$\begin{array}{cccccccc} \begin{bmatrix} 4 \\ 1 \\ 2 \\ 3 \end{bmatrix} & \xrightarrow{t_{\uparrow 3}} & \begin{bmatrix} 2 \\ 4 \\ 1 \\ 3 \end{bmatrix} & \xrightarrow{t_{\uparrow 4}} & \begin{bmatrix} 3 \\ 2 \\ 4 \\ 1 \end{bmatrix} & \xrightarrow{t_{\uparrow 4}} & \begin{bmatrix} 1 \\ 3 \\ 2 \\ 4 \end{bmatrix} & \xrightarrow{t_{\uparrow 4}} & \begin{bmatrix} 4 \\ 1 \\ 3 \\ 2 \end{bmatrix} & \xrightarrow{t_{\uparrow 3}} & \begin{bmatrix} 3 \\ 4 \\ 1 \\ 2 \end{bmatrix} & \xrightarrow{t_{\uparrow 4}} & \begin{bmatrix} 2 \\ 3 \\ 4 \\ 1 \end{bmatrix} & \xrightarrow{t_{\uparrow 4}} & \begin{bmatrix} 1 \\ 2 \\ 3 \\ 4 \end{bmatrix} \\ \underbrace{\phantom{\begin{bmatrix} 4 \\ 1 \\ 2 \\ 3 \end{bmatrix}}}_{C''} & & & & & & & & & & & & & & \end{array}$$

Finally, we observe that shifting the resulting code so it begins with Id is satisfactory. □

We remark that, while Theorem 8 does not produce auxiliary codes much larger than the parity-preserving code of size $\frac{|S_{2m}|}{2m}$, it does at least allow us to permute the last element.

Next, we present another construction which yields larger codes, for even $k \geq 6$ (but not $k = 4$). From now on, we fix $m \geq 2$. We also define $\varphi : S_{2m+2} \rightarrow S_{2m+2}$ by

$$\varphi = t_{\uparrow 2m+2}^2 \circ t_{\uparrow 2m-1}^{-1}.$$

We note that

$$\varphi(\pi) = \pi \circ (1, 2m+1)(2m+2, 2m, 2m-1, \dots, 2),$$

Hence, informally, in π 's vector notation, φ transposes the elements in indices $1, 2m+1$, and cyclically shifts all other elements once to the top. We can also observe that $\varphi^{2m} = \text{Id}$.

We conveniently define, for all $r \geq 0$, the permutations

$$\begin{aligned} \hat{\pi}_r &= \varphi^r(\text{Id}) \\ &= (1, 2m+1)^r (2m+2, 2m, 2m-1, \dots, 2)^r \in S_{2m+2}, \end{aligned}$$

In particular, we note that when $r \equiv r' \pmod{2m}$, and only then, we have $\hat{\pi}_r = \hat{\pi}_{r'}$.

Lemma 10 For all $r \geq 0$ a parity-preserving $G_{\uparrow}(2m+2, M_{2m+2})$ code P_r exists which begins with $\hat{\pi}_r$ and ends with $t_{\uparrow 2m-1}^{-1} \hat{\pi}_r$, where

$$M_{2m+2} = \begin{cases} 57 & m = 2, \\ \frac{(2m+1)!}{2} & m > 2. \end{cases}$$

Proof: The claim follows trivially from Lemma 6, if we shift the generating transition sequence such that it ends with $t_{\uparrow 2m-1}$ and apply it to $\hat{\pi}_r$, due to Lemma 5. ■

We note in particular that for all r , $\hat{\pi}_r$ is even, and thus $P_r \subseteq A_{2m+2}$. Moreover, since the parity-preserving code P_r does not employ $t_{\uparrow 2m+2}$, for all $\pi \in P_r$ it holds that

$$\begin{aligned} \pi(2m+2) &= \hat{\pi}_r(2m+2) \\ &= \begin{cases} 2m+2 & r \equiv 0 \pmod{2m}, \\ 2m+1 - (r \bmod 2m) & r \not\equiv 0 \pmod{2m}. \end{cases} \end{aligned}$$

Thus, when considered as sets,

$$P_r \cap P_{r'} = \emptyset,$$

for all $0 \leq r < r' < 2m$.

We shall construct a $G_{\uparrow}^{\text{aux}}(2m+2, M)$ code by stitching together $P_1, P_2, \dots, P_{2m-1}$. We will need to amend P_0 before incorporating it into our code, for reasons we shall discuss below. First, we describe the stitching method in the following lemma.

Lemma 11 For all $r \geq 0$ (including, in particular, $r = 2m-1$), we may concatenate P_r, P_{r+1} into a (non-cyclic) ‘‘push-to-the-top’’ code by applying the transitions $t_{\uparrow 2m+2}, t_{\uparrow 2m+2}$ to the last permutation of P_r , which is $t_{\uparrow 2m-1}^{-1} \hat{\pi}_r$. Additionally, the only odd permutation in the resulting code is

$$\beta_{r+1} = t_{\uparrow 2m+2}^{-1}(\hat{\pi}_{r+1}).$$

We shall refer to it as the $(r+1)$ -bridge.

Proof: The claim follows trivially from the definition

$$\hat{\pi}_{r+1} = \varphi(\hat{\pi}_r) = t_{\uparrow 2m+2} \circ (t_{\uparrow 2m+2} \circ t_{\uparrow 2m-1}^{-1}(\hat{\pi}_r)),$$

since P_r, P_{r+1} are parity-preserving, and $t_{\uparrow 2m+2}$ flips parity. ■

Lemma 11 can be used iteratively to concatenate $P_1, P_2, \dots, P_{2m-1}$, with a single odd permutation—the r -bridge—between each pair of P_{r-1}, P_r . Thus, we obtain the sequence

$$P_1, \beta_2, P_2, \beta_3, \dots, \beta_{2m-1}, P_{2m-1}.$$

Note that if any two permutations π_1, π_2 in the resulting sequence satisfy $\pi_1 = (q, 2m+2) \circ \pi_2$ for some $q \in [2m+1]$, then w.l.o.g π_2 is odd and hence an r -bridge for some r , and π_1 is even and thus not a bridge. Since in every bridge the last element is

$$\beta_r(2m+2) = \hat{\pi}_r(1) \in \{1, 2m+1\},$$

and in every non-bridge it is not, it must follow, then, that $q = \beta_r(2m+2)$, and in particular

$$\begin{aligned} \pi_1(2m+2) &= (\beta_r(2m+2), 2m+2) \circ \beta_r(2m+2) \\ &= 2m+2, \end{aligned}$$

thus $\pi_1 \in P_0$.

We witness, therefore, that no such pair of permutations exist, since we have not yet incorporated P_0 into our code. It also becomes apparent that P_0 must necessarily be amended prior to its inclusion, so it does not include any permutations of the form

$$(\beta_r(2m+2), 2m+2) \circ \beta_r, \quad 0 < r \leq 2m.$$

In order to do so, we note that for all $r \geq 0$

$$\beta_r(2m) = \hat{\pi}_r(2m+1) \in \{1, 2m+1\},$$

and in particular $\beta_r(2m) \neq 2m+2$, hence

$$(\beta_r(2m+2), 2m+2) \circ \beta_r(2m) = \beta_r(2m) \in \{1, 2m+1\}.$$

It follows that if we let P'_0 be generated by the transition sequence $t_{\uparrow 2m-1}^{2m-1}$ applied to $\hat{\pi}_0$, then it is parity-preserving, its last permutation is $t_{\uparrow 2m-1}^{-1} \hat{\pi}_0$, and for all $\pi \in P'_0$ we have $\pi(2m) = 2m \notin \{1, 2m+1\}$, thus

$$P'_0 \cap \{(\beta_r(2m+2), 2m+2) \circ \beta_r\}_{r=1}^{2m} = \emptyset.$$

Lemma 12 The following sequence P ,

$$P = P'_0, \beta_1, P_1, \beta_2, P_2, \beta_3, \dots, \beta_{2m-1}, P_{2m-1}, \beta_{2m},$$

is a cyclic $G_{\uparrow}^{\text{aux}}(2m+2, M)$.

Proof: By Lemma 11, and since when considered as sets,

$$P_r \cap P_{r'} = \emptyset$$

for all $0 < r < r' < 2m$, and similarly P'_0 is disjoint from $P_1, P_2, \dots, P_{2m-1}$, we know that P is a $G_{\uparrow}(2m+2, M)$.

As seen above, if for any two permutations $\pi_1, \pi_2 \in P$ and $q \in [2m+1]$ we have $\pi_1 = (q, 2m+2) \circ \pi_2$, then w.l.o.g. $\pi_2 = \beta_r$ for some $0 < r \leq 2m$ and $q = \beta_r(2m+2)$. In particular, $\pi_1(2m+2) = 2m+2$, thus

$$\pi_1 \notin \{\beta_r\}_{r=1}^{2m} \cup \bigcup_{r=1}^{2m-1} P_r,$$

thus $\pi_1 \in P'_0$. But

$$P'_0 \cap \{(\beta_r(2m+2), 2m+2) \circ \beta_r\}_{r=1}^{2m} = \emptyset,$$

in contradiction. \blacksquare

The auxiliary code from Lemma 12 is almost what we need. The only property lacking is the fact that Id is not followed in P by the transition $t_{\uparrow 2m+2}$. We fix this in the following theorem.

Theorem 13 Let $k \geq 6$ be even. Then there exists a $G_{\uparrow}^{\text{aux}}(k, M)$ starting at Id and a $t_{\uparrow k}$ transition, with

$$M = \begin{cases} 178 & k = 6, \\ (k-3) \left(\frac{(k-1)!}{2} + 2 \right) + 1 & k > 6. \end{cases}$$

In particular, for all $k > 6$,

$$M > \frac{k-3}{k} \cdot \frac{k!}{2}.$$

Proof: Denote $k = 2m+2$ for $m \geq 2$, and let $P = (c_j)_{j=1}^M$ be the code from Lemma 12. Since Id $\in S_k$ is not followed with a $t_{\uparrow k}$ transition in P , we denote the last permutation of P'_0 by $\tilde{\pi}$, and replace P with

$$\tilde{P} = \tilde{\pi}^{-1} P = \left(\tilde{\pi}^{-1} \circ c_j \right)_{j=1}^M.$$

We observe that \tilde{P} is still a ‘‘push-to-the-top’’ code since ‘‘push-to-the-top’’ transitions are group actions by right-multiplications. Moreover, since $\tilde{\pi}(k) = k$, if for some $\pi_1, \pi_2 \in P$ we have $\tilde{\pi}^{-1} \circ \pi_1 = (q, k) \circ (\tilde{\pi}^{-1} \circ \pi_2)$, where $q \in [k-1]$, then

$$\begin{aligned} \pi_1 &= \tilde{\pi} \circ \left[(q, k) \circ (\tilde{\pi}^{-1} \circ \pi_2) \right] \\ &= \left[\tilde{\pi} \circ (q, k) \circ \tilde{\pi}^{-1} \right] \circ \pi_2 = (\tilde{\pi}(q), k) \circ \pi_2, \end{aligned}$$

and $\tilde{\pi}(q) \in [k-1]$, in contradiction.

As for the size of the code, note that $|P'_0| = 2m-1 = k-3$ and

$$|P_1| = |P_2| = \dots = |P_{2m-1}| = \begin{cases} 57 & k = 6, \\ \frac{(k-1)!}{2} & k > 6. \end{cases}$$

Counting $\beta_1, \dots, \beta_{2m}$, the claim is thus substantiated. \blacksquare

To conclude this section, we combine Lemma 6, Theorem 8 and Theorem 13 into the following corollary.

Corollary 14 For all $k \geq 3$ there exists a $G_{\uparrow}^{\text{aux}}(k, \tilde{M}_k)$ starting with Id and a $t_{\uparrow k}$ transition, where

$$\tilde{M}_k = \begin{cases} 8 & k = 4; \\ 57 & k = 5; \\ 178 & k = 6; \\ \frac{k!}{2} & 5 \neq k \equiv 1 \pmod{2}; \\ \rho_k \frac{k!}{2} & 6 < k \equiv 0 \pmod{2}, \end{cases}$$

where $\rho_k > \frac{k-3}{k}$.

IV. CODE CONSTRUCTION

In this section we present the main construction of our paper, and discuss the size and asymptotic rate of the resulting codes. We will show, surprisingly, that our method generates codes which are larger than formerly known families of codes, even though we require the additional structure of a Gray code.

A. Main code construction

We now present a construction of $G_{\uparrow}(n, M, d)$ codes, for $d \leq n$, which we base on Corollary 14 and Lemma 7.

It will simplify the presentation to assume $n = kd$ for some positive $k \geq 2$, since in that case every congruence class modulo d of $[n]$ has size k . Nonetheless, the construction is applicable to any $n > d$ with natural amendments. We discuss these changes, focusing on special cases, after presenting the simple construction first.

Construction A Let $n, k, d \in \mathbb{N}$, with $n = kd$ and $k \geq 2$. We recursively construct a sequence of codes, C_d, C_{d-1}, \dots, C_1 . An explicit construction is given for C_d and a recursion step constructs C_m from C_{m+1} .

Recursion base: We construct the code C_d by starting at the permutation $\sigma_0 \in S_n$ defined by

$$\sigma_0(j) = d(j \bmod k) + \left\lceil \frac{j}{k} \right\rceil.$$

We obtain a transition sequence $t_{\uparrow r_1}, t_{\uparrow r_2}, \dots, t_{\uparrow r_k!}$ which generates the $G_{\uparrow}(k, k!)$ provided by Lemma 7. The code C_d starts with σ_0 , and uses the transition sequence

$$t_{k(d-1)+1\uparrow k(d-1)+r_1}, t_{k(d-1)+1\uparrow k(d-1)+r_2}, \dots \\ \dots, t_{k(d-1)+1\uparrow k(d-1)+r_k!}.$$

Recursion step: Assume C_{m+1} has already been constructed, starting with permutation σ_0 . Additionally, let

$$t_{\uparrow k+1}, t_{\uparrow i_2}, \dots, t_{\uparrow i_{\tilde{M}_{k+1}}} \quad (1)$$

be a transition sequence generating a $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code provided by Corollary 14.

We construct the code C_m as follows: replace each $t_{km+1\uparrow j}$ transition of C_{m+1} with $t_{k(m-1)+1\uparrow j}$, followed by $t_{k(m-1)+1\uparrow k(m-1)+i_2}$, $t_{k(m-1)+1\uparrow k(m-1)+i_3}$, and so on until $t_{k(m-1)+1\uparrow k(m-1)+i_{\tilde{M}_{k+1}}}$. \square

Lemma 15 For all $n = kd$, $k \geq 2$, the code C_d from Construction A is a $G_{k(d-1)+1\uparrow}(n, k!, d)$.

Proof: The parameters of the code are obvious, except perhaps the minimal distance d . The fact that the codewords of C_d are distinct follows from Lemma 7.

To prove the minimal distance d , note that for all $0 \leq u < d$ and $ku + 1 \leq i < j \leq k(u+1)$ it holds that $\sigma_0(i) \equiv \sigma_0(j) \pmod{d}$. Thus, for every distinct $\sigma, \tau \in C_d$, there exists j , $m(d-1) < j \leq kd = n$, such that $\sigma(j) \not\equiv \tau(j)$. Since by construction $\sigma(j) \equiv \tau(j) \equiv 0 \pmod{d}$, we observe

$$d_{\infty}(\sigma, \tau) \geq |\sigma(j) - \tau(j)| \geq d,$$

implying that C_d is a $G_{k(d-1)+1\uparrow}(n, k!, d)$. \blacksquare

Example 16 We let $d = 3$, $k = 2$, and $n = kd = 6$. We construct the code C_3 starting at

$$\sigma_0 = [4, 1, 5, 2, 6, 3] \in S_6.$$

We use the complete $G_{\uparrow}(2, 2)$ shown in Example 9, which is generated by the sequence $t_{\uparrow 2}, t_{\uparrow 2}$. We arrive at a generating sequence $t_{5\uparrow 6}, t_{5\uparrow 6}$ for C_3 . Hence, in our example

$$C_3 = ([4, 1, 5, 2, 6, 3], [4, 1, 5, 2, 3, 6]),$$

which is readily seen to be a $G_{5\uparrow}(6, 2, 3)$ code. \square

Theorem 17 For all $n = kd$, $k \geq 2$, the code C_1 from Construction A is a $G_{\uparrow}(n, \tilde{M}_{k+1}^{d-1} \cdot k!, d)$.

Proof: To prove the claim we will prove by induction that C_m from Construction A, for all $m \in [d]$, is a $G_{k(m-1)+1\uparrow}(n, \tilde{M}_{k+1}^{d-m} \cdot k!, d)$. The base case of C_d was proved in Lemma 15. Assume the claim holds for C_{m+1} and we now prove it for C_m .

Recall (1) gives the sequence of transitions for a $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$. Then

$$t_{\uparrow i_{\tilde{M}_{k+1}}} t_{\uparrow i_{\tilde{M}_{k+1}-1}} \dots t_{\uparrow i_3} t_{\uparrow i_2} = t_{\uparrow k+1}^{-1}.$$

Thus,

$$t_{km+1\uparrow j} = \left(\prod_{r=2}^{\tilde{M}_{k+1}} t_{k(m-1)+1\uparrow k(m-1)+i_r} \right) t_{k(m-1)+1\uparrow j}$$

(where the product is expanded right-to-left). Therefore, C_m expands each “push-to-the- $km+1$ st-index” transition of C_{m+1} into \tilde{M}_{k+1} “push-to-the- $k(m-1)+1$ st-index” transitions.

It follows that C_m contains the codewords of C_{m+1} in the same order, with $\tilde{M}_{k+1} - 1$ new words inserted between any two words originally from C_{m+1} . We say that each codeword of C_{m+1} (now appearing in C_m) is the C_{m+1} -parent of each of the \tilde{M}_{k+1} preceding codewords in C_m (including itself), since their vector notations agree on the order of the elements

$$\sigma_0(km+1), \sigma_0(km+2), \dots, \sigma_0(n).$$

Now, suppose that $\sigma, \tau \in C_m$ satisfy $d_{\infty}(\sigma, \tau) < d$. Let σ', τ' be their C_{m+1} -parents, respectively. To complete the proof we will show that $\sigma = \tau$.

Case 1: $\sigma' = \tau'$. Denote

$$x = \sigma'(km+1) = \tau'(km+1)$$

and $s = \sigma^{-1}(x)$, $a = \tau(s)$.

If $a = x$ then for all $j \neq s$, $k(m-1) < j \leq km+1$, we have $\sigma(j) \equiv \tau(j) \pmod{d}$ and

$$|\sigma(j) - \tau(j)| \leq d_{\infty}(\sigma, \tau) < d,$$

hence $\sigma(j) = \tau(j)$, and $\sigma = \tau$.

Otherwise, $a \neq x$, and denote $t = \tau^{-1}(x) \neq s$. It similarly holds for all $j \notin \{s, t\}$, $k(m-1) < j \leq km+1$, that $\sigma(j) = \tau(j)$. We therefore observe $\tau = \sigma \circ (s, t)$. This implies that, if we let $\hat{\sigma}, \hat{\tau} \in S_{k+1}$ be the permutations in the $G_{\uparrow}^{\text{aux}}$ we obtained, generated similarly to σ, τ , respectively (i.e., by their corresponding transition sequences), then

$$\hat{\tau} = \hat{\sigma} \circ (s - k(m-1), t - k(m-1)) = (q, k+1)\hat{\sigma}$$

for some $q \in [k]$, in contradiction to the fact it was a $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$.

Case 2: $\sigma' \neq \tau'$. Since $\sigma', \tau' \in C_{m+1}$ we have by assumption $d_{\infty}(\sigma', \tau') \geq d$, and note that for all j satisfying $j \leq k(m-1)$ or $j > km+1$, it holds that $\sigma(j) = \sigma'(j)$ and $\tau(j) = \tau'(j)$. Hence there exists j , $k(m-1) < j \leq km+1$, such that

$$|\sigma(j) - \tau(j)| < d \quad \text{but} \quad |\sigma'(j) - \tau'(j)| \geq d.$$

Note particularly, since for all $k(m-1) < j \leq km$ it holds that $\sigma'(j) = \sigma_0(j) = \tau'(j)$, that we have

$$|\sigma'(km+1) - \tau'(km+1)| \geq d.$$

Denote $x = \sigma'(km+1)$, $y = \tau'(km+1)$, and note that

$$\{\sigma(j)\}_{j=k(m-1)+1}^{km+1} = \{a_i\}_{i=1}^k \cup \{x\};$$

$$\{\tau(j)\}_{j=k(m-1)+1}^{km+1} = \{a_i\}_{i=1}^k \cup \{y\},$$

where $\{a_i\}_{i=1}^k$ is a congruence class modulo d of $[n]$, of which x, y are not members.

Let $s = \sigma^{-1}(x)$ and denote $a = \tau(s)$. Since

$$|x - a| = |\sigma(s) - \tau(s)| \leq d_\infty(\sigma, \tau) < d$$

we have $a \neq y$. Let $t = \sigma^{-1}(a)$. Since $a \in \{a_i\}_{i=1}^k$ is a congruence class modulo d , for all $b \in \{a_i\}_{i=1}^k \setminus \{a\}$ we observe $|a - b| \geq d$, but

$$|a - \tau(t)| = |\sigma(t) - \tau(t)| \leq d_\infty(\sigma, \tau) < d$$

and therefore $\tau(t) = y$. For all $j \notin \{s, t\}$ satisfying $k(m-1) < j \leq km+1$ we then have $\sigma(j) \equiv \tau(j) \pmod{d}$ and $|\sigma(j) - \tau(j)| \leq d_\infty(\sigma, \tau) < d$, hence $\sigma(j) = \tau(j)$.

This implies that, if we again let $\hat{\sigma}, \hat{\tau} \in S_{k+1}$ be the permutations in the $G_{\uparrow}^{\text{aux}}$ generated similarly to σ, τ respectively, then

$$\hat{\tau} = \hat{\sigma} \circ (s - k(m-1), t - k(m-1)) = (q, k+1)\hat{\sigma}$$

where q is given by $a = a_q \in \{a_i\}_{i=1}^k$, again contradicting the properties of a $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$. Hence C_m has minimal ℓ_∞ -distance of at least d , as required. ■

Example 18 We complete Example 16 into a $G_{\uparrow}(6, 3^2 \cdot 2, 3)$ code by applying the recursion step twice. In each step, since $k=2$, we utilize the trivial parity-preserving $G_{\uparrow}^{\text{aux}}(3, 3)$ code generated by the sequence $t_{\uparrow 3}, t_{\uparrow 3}, t_{\uparrow 3}$.

Firstly, recall that we used

$$\sigma_0 = [4, 1, 5, 2, 6, 3] \in S_6,$$

and the sequence $t_{5\uparrow 6}, t_{5\uparrow 6}$ generates

$$C_3 = ([4, 1, 5, 2, 6, 3], [4, 1, 5, 2, 3, 6]).$$

We build C_2 by exchanging each $t_{5\uparrow 6}$ transition by $t_{3\uparrow 6}$ followed by 2 instances of $t_{2+1\uparrow 2+3} = t_{3\uparrow 5}$; the middle level of Figure 1 shows the resulting code.

Secondly, as seen in the same figure, each $t_{3\uparrow j}$ transition of C_2 , $j \in \{5, 6\}$, can be replaced by $t_{1\uparrow j} = t_{\uparrow j}$, followed by 2 instances of $t_{0+1\uparrow 0+3} = t_{\uparrow 3}$, to generate C_1 .

Note that $C_3 \subseteq C_2 \subseteq C_1$, and that they are $G_{5\uparrow}(6, 2, 3)$, $G_{3\uparrow}(6, 6, 3)$ and $G_{\uparrow}(6, 18, 3)$ codes, respectively. □

We now describe the changes needed in Construction A to allow general n and d parameters. We first consider n not necessarily being a multiple of d , but still $n \geq 2d$. For all $i \in [d]$, let

$$\mathcal{R}_i = \{i, i+d, i+2d, \dots, n - ((n-i) \bmod d)\},$$

be the i th congruence class modulo d of $[n]$. Then

$$|\mathcal{R}_i| = \begin{cases} \lfloor \frac{n}{d} \rfloor & 1 \leq i \leq (n \bmod d), \\ \lfloor \frac{n}{d} \rfloor + 1 & (n \bmod d) < i \leq d. \end{cases}$$

We define the starting permutation

$$\sigma_0 = [\mathcal{R}_1 | \mathcal{R}_2 | \dots | \mathcal{R}_d] \in S_n,$$

to be comprised of a concatenation of the congruence classes, where the order of elements within the congruence class is arbitrary. Additionally, the recursion base uses a $G_{\uparrow}(|\mathcal{R}_d|, |\mathcal{R}_d|!)$. As for the recursion step of constructing C_m from C_{m+1} , we can still apply it with the following changes:

- We choose $G_{\uparrow}^{\text{aux}}(|\mathcal{R}_m| + 1, \tilde{M}_{|\mathcal{R}_m|+1})$.
- We use push operations to position $1 + \sum_{i=1}^{m-1} |\mathcal{R}_i|$.

We obtain C_1 which is a $G_{\uparrow}(n, M, d)$, where

$$M = \tilde{M}_{\lfloor n/d \rfloor + 1}^{n \bmod d} \cdot \tilde{M}_{\lfloor n/d \rfloor + 1}^{d - (n \bmod d) - 1} \cdot \left\lfloor \frac{n}{d} \right\rfloor !.$$

Finally, we discuss the special case of $n < 2d$, in which all but $(n \bmod d)$ congruence classes are singletons. We will amend our construction by replacing the recursion base with

$$C_m = \left\{ \sigma_0, t_{2m-1\uparrow 2m+1} \sigma_0, t_{2m-1\uparrow 2m+1}^2 \sigma_0 \right\},$$

where $m = n \bmod d$, and continuing the recursion step as discussed above. Thus, we are effectively only using the first member of \mathcal{R}_{m+1} together with the previous congruence classes, fixing $\sigma_0(j)$ for $j > 2m+1$. In this case, we obtain C_1 which is a $G_{\uparrow}(n, 3^{n \bmod d}, d)$.

Thus, in what follows, whenever we mention Construction A, we refer to its most general version applying to all n and d .

B. Code-size analysis and comparison

We would like to give an explicit expression for the size of the codes constructed by Construction A. This would enable a comparison with previously known results.

Lemma 19 Let C_1 be the code from (the general version of) Construction A. Then its size, $|C_1|$, is given by (2).

Proof: Let us first assume $n \geq 2d$. We note the asymmetry in Construction A between congruence classes \mathcal{R}_i of odd and even sizes. Indeed, a class of size $|\mathcal{R}_i| = k \geq 2$ (for all classes other than \mathcal{R}_d , which is used in the recursion base and whose contribution is based on the $G_{\uparrow}(k, k!)$ code) contributes to the code size, according to Corollary 14, a multiplicative factor of

$$\tilde{M}_{k+1} = \begin{cases} 8 & k = 3; \\ 57 & k = 4; \\ 178 & k = 5; \\ \frac{(k+1)!}{2} & 4 \neq k \equiv 0 \pmod{2}; \\ \rho_{k+1} \frac{(k+1)!}{2} & 5 < k \equiv 1 \pmod{2}, \end{cases}$$

where, again, $\rho_{k+1} > \frac{k-2}{k+1}$.

It is therefore important to note that when $\lfloor \frac{n}{d} \rfloor \equiv 0 \pmod{2}$, $[n]$ has $(n \bmod d)$ congruence classes modulo d of odd size $\lfloor \frac{n}{d} \rfloor$, and $d - (n \bmod d)$ classes of even size $\lfloor \frac{n}{d} \rfloor$. Thus, if additionally $\lfloor \frac{n}{d} \rfloor > 4$, the constructed code C_1 is of size

$$|C_1| = \left(\rho_{\lfloor n/d \rfloor + 1} \frac{(\lfloor n/d \rfloor + 1)!}{2} \right)^{n \bmod d} \cdot \left\lfloor \frac{n}{d} \right\rfloor ! \cdot \left(\frac{(\lfloor n/d \rfloor + 1)!}{2} \right)^{d - (n \bmod d) - 1},$$

and simple rearranging gives us the first case of (2). Similar considerations give us the next five cases of (2).

Finally, we consider the case of $n < 2d$, which implies $\lfloor \frac{n}{d} \rfloor = 1$. In this special case we only permute $(n \bmod d) =$

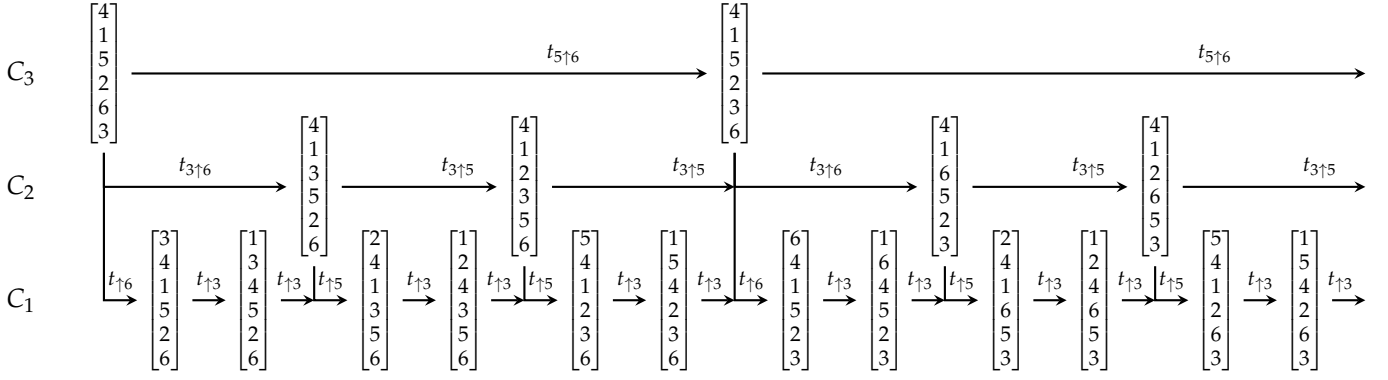


Figure 1. Construction A as demonstrated in the case $d = 3, k = 2$.

$$|C_1| = \begin{cases} \left(\left\lfloor \frac{n}{d} \right\rfloor + 1 \right)^{n \bmod d} \left(\left\lfloor \frac{n}{d} \right\rfloor + 1 \right)! \cdot \frac{\rho_{\lfloor n/d \rfloor + 1}^{n \bmod d}}{2^{d-1} (\lfloor n/d \rfloor + 1)} & 4 < \left\lfloor \frac{n}{d} \right\rfloor \equiv 0 \pmod{2}, \\ \left(\frac{178}{57} \right)^{n \bmod d} \cdot 57^{d-1} \cdot 24 & \left\lfloor \frac{n}{d} \right\rfloor = 4, \\ \left(\frac{8}{3} \right)^{n \bmod d} \cdot 3^{d-1} \cdot 2 & \left\lfloor \frac{n}{d} \right\rfloor = 2, \\ \left(\left\lfloor \frac{n}{d} \right\rfloor + 1 \right)^{n \bmod d} \left(\left\lfloor \frac{n}{d} \right\rfloor + 1 \right)! \cdot \frac{\rho_{\lfloor n/d \rfloor + 1}^{(d-1) - (n \bmod d)}}{2^{d-1} (\lfloor n/d \rfloor + 1)} & 5 < \left\lfloor \frac{n}{d} \right\rfloor \equiv 1 \pmod{2}, \\ \left(\frac{1260}{89} \right)^{n \bmod d} \cdot (178)^d \cdot \frac{120}{178} & \left\lfloor \frac{n}{d} \right\rfloor = 5, \\ \left(\frac{57}{8} \right)^{n \bmod d} \cdot 8^d \cdot \frac{3}{4} & \left\lfloor \frac{n}{d} \right\rfloor = 3, \\ 3^{n \bmod d} & \left\lfloor \frac{n}{d} \right\rfloor = 1. \end{cases} \quad (2)$$

$(n-d)$ congruence classes of $[n]$, (and each such class has $2 = \lfloor \frac{n}{d} \rfloor + 1$ elements). As mentioned, we therefore construct a code of size $|C_1| = 3^{n \bmod d}$. ■

We comment that it is also possible to achieve a slight gain in code size by reordering σ_0 so that the last block consists of a congruence class of odd size, rather than even, where the added complexity of index calculation is inconsequential. The asymptotic gain in code rate vanishes.

We now turn to comparing the size of the resulting code with that of previously constructed codes, as well as known bounds on the cardinality of such codes.

The first comparison we make is with codes that have the Gray property. Such codes were only studied for $d = 2$, i.e., snake-in-the-box codes or $G_{\uparrow}(n, M, 2)$ codes in our notation. These codes were studied in [36, Thm. 24], where it was shown that such codes can be constructed with sizes

$$M = \left\lfloor \frac{n}{2} \right\rfloor! \left(\left\lfloor \frac{n}{2} \right\rfloor + \left(\left\lfloor \frac{n}{2} \right\rfloor - 1 \right)! \right).$$

Construction A improves this size by a factor of $\frac{1}{2} \left(\left\lfloor \frac{n}{2} \right\rfloor + 1 \right) \left\lfloor \frac{n}{2} \right\rfloor$, times $\rho_{\lfloor n/2 \rfloor + 1}$ when $n \equiv 2 \pmod{4}$ (in the case of $n \equiv 1 \pmod{4}$) $\rho_{\lfloor n/2 \rfloor + 1}$ is eliminated by changing the order of congruence classes in σ_0). We note that a similar improvement was made concurrently by [35] in a preprint devoted solely to the case of $d = 2$, i.e., snake-in-the-box codes.

We now also compare our results to error-correcting codes with the ℓ_{∞} -metric which are not necessarily Gray codes (LMRM-codes). We observe that the best known general

LMRM-code construction to date, [32, Cst. 1, Thm. 2] and [24, Sec. III-A], presented (n, M, d) -LMRM codes with sizes

$$M = \left\lfloor \frac{n}{d} \right\rfloor!^{n \bmod d} \left\lfloor \frac{n}{d} \right\rfloor!^{d - (n \bmod d)},$$

which our construction improves upon, more pronouncedly the more $[n]$ has even-sized congruence classes modulo d (cf. (2)).

In the asymptotic regime, we go on to examine the case of $d = \Theta(n)$. For an (n, M, d) -LMRM code (and in particular a $G_{\uparrow}(n, M, d)$), we follow the convention (e.g., [32]) of defining the *rate* of the code

$$R = \frac{1}{n} \log_2 M,$$

and its *normalized distance*

$$\delta = \frac{d}{n}.$$

The following were proven in [32].

Lemma 20 [32, Thm. 23] For any (n, M, d) -LMRM code it holds that

$$R \leq 2 - 2\delta \left\lfloor \frac{1}{\delta} \right\rfloor - \left(\delta \left\lfloor \frac{1}{\delta} \right\rfloor - \delta \right) \log_2(\delta) - \left(1 + \delta - \delta \left\lfloor \frac{1}{\delta} \right\rfloor \right) \log_2 \left(1 + \delta - \delta \left\lfloor \frac{1}{\delta} \right\rfloor \right) + o(1).$$

Lemma 21 [32, Thm. 27] For any $0 < \delta \leq 1$ the construction of [32, Cst. 1, Thm. 2] and [24, Sec. III-A] yields codes with

$$R = \left(1 - \delta \left\lfloor \frac{1}{\delta} \right\rfloor\right) \log_2 \left(\left\lfloor \frac{1}{\delta} \right\rfloor! \right) + \left(\delta + \delta \left\lfloor \frac{1}{\delta} \right\rfloor - 1 \right) \log_2 \left(\left\lfloor \frac{1}{\delta} \right\rfloor! \right).$$

It was also shown in [32] that a Gilbert-Varshamov-like non-constructive bound exists:

Lemma 22 [32, Thm. 25] For any $0 < \delta \leq 1$ there exist (n, M, d) -LMRM codes satisfying $\frac{d}{n} \geq \delta$ with rate $R \geq f_{\text{GV}}(\delta) + o(1)$, where

$$f_{\text{GV}} = \begin{cases} \log_2 \frac{1}{\delta} + 2\delta (\log_2 e - 1) - 1 & 0 < \delta \leq \frac{1}{2} \\ -2\delta \log_2 \frac{1}{\delta} + 2(1 - \delta) \log_2 e & \frac{1}{2} \leq \delta \leq 1 \end{cases}$$

We therefore aim to show that our construction can bridge some of the gap between the given bounds and known constructions.

Lemma 23 Let C_1 be the code from (the general version of) Construction A. Then an estimate from below of its rate R as a function of its normalized distance δ is given by (3).

Proof: The proof follows by a simple substitution of $(n \bmod d) = n - d \lfloor \frac{n}{d} \rfloor$ and $d = n\delta$ into (2). We also recall that $\rho_k > \frac{k-3}{k}$. ■

In conclusion, these asymptotic rates and bounds are shown in Figure 2. We note in particular that the rate of codes produced by Construction A is strictly higher than that of previously known constructions (as in Lemma 21). Furthermore, it produces codes with rates higher than those guaranteed by the Gilbert-Varshamov-like Lemma 22 for all δ greater than ≈ 0.1 except in a small neighborhood of $\frac{1}{5}$, whereas known constructions only bypassed these rates only for δ greater than ≈ 0.349 .

V. DECODING ALGORITHM

This section is devoted to devising a decoding algorithm capable of correcting a noisy received version of a transmitted codeword.

Known constructions of (n, M, d) -LMRM codes, presented in [32, Cst. 1, Thm. 2] and [24, Sec. III-A], lend themselves to straightforward decoding algorithms, efficiently done in $O(n)$ operations, since for any given codeword σ and index $i \in [n]$, $r = \lfloor \sigma(i) \bmod d \rfloor$ is known. Hence, if a retrieved permutation τ satisfies $d_\infty(\sigma, \tau) \leq \lfloor (d-1)/2 \rfloor$, then $\sigma(i)$ is known to be the unique element $k \in r + d\mathbb{Z}$ satisfying $|k - \tau(i)| \leq \lfloor (d-1)/2 \rfloor$.

Our proposed construction diverges from that rigid partition. However, we can still efficiently decode noisy information, provided errors of magnitude no more than t have occurred, where $2t + 1 \leq d$. More precisely, we assume that for every stored permutation σ and retrieved permutation τ it holds that $d_\infty(\sigma, \tau) \leq t \leq \lfloor (d-1)/2 \rfloor$.

To simplify our presentation we assume $n = kd$, since then our construction only makes (repeated) use of a single

auxiliary $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code. Extensions to the general version of Construction A are easily obtainable.

We first require a function `ValidAux` capable of detecting whether a given permutation $\sigma \in S_{k+1}$ belongs to the auxiliary $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code.

Lemma 24 For an auxiliary $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code provided by Lemma 6, Theorem 8 or Theorem 13, a function `ValidAux` can be implemented to operate in $O(k)$ steps.

Proof: If we use Lemma 6, then the auxiliary code consists of all even permutations, and it is well known that we can determine the signature of a permutation $\sigma \in S_{k+1}$ in $O(k)$ operations, e.g., by finding a cycle decomposition of σ .

If we instead use Theorem 8, then the auxiliary code consists of exactly those permutations in whose vector notation 1 follows $k+1$ (cyclically), i.e.,

$$\sigma \left((\sigma^{-1}(k+1) \bmod (k+1)) + 1 \right) = 1,$$

which again requires $O(k)$ steps to verify.

Finally, for Theorem 13 we divide into cases according to $\sigma(k)$. For all elements other than $1, k-1, k$ the problem again reduces to determining sign σ . For those elements, only cyclic shift (on a subset of indices, by case) of a known permutation are valid, which we can easily verify in linear time. ■

An important notion of a *window* will be useful. Let $\sigma \in S_n$ be a permutation, $n = kd$. For all $j \in [d]$ we define the j th window as the set of indices

$$W_j = \{k(j-1) + 2, k(j-1) + 3, \dots, kj + 1\} \cap [n].$$

The windows partition $[n] \setminus \{1\}$, and are all of size k except W_d which is of size $k-1$.

Given a set $I \subseteq [n]$, we conveniently denote

$$\sigma(I) = \{\sigma(i) \mid i \in I\}.$$

We prove a simple lemma concerning properties of windows of codewords from Construction A.

Lemma 25 Let σ be a codeword of C_1 from Construction A, with $n = kd$. Then for all $j \in [d]$,

$$k-1 \leq |\sigma(W_j) \cap \mathcal{R}_j| \leq k,$$

i.e., at most one element of $\sigma(W_j)$ does not leave a residue of j modulo d . In particular, $\sigma(W_d) \subseteq \mathcal{R}_d$.

Additionally, if $|\sigma(W_j) \cap \mathcal{R}_j| = k-1$, $j \in [d-1]$, then $x \in \sigma(W_j) \setminus \mathcal{R}_j$ satisfies $x \in \mathcal{R}_{j'}$ for some $j' > j$.

Proof: Take any $1 < j \in [d]$, and let σ_j be the C_j -parent of σ . Then, in C_1 , no transition between σ and σ_j is induced by C_j , and hence σ_j is derived from σ by a (perhaps empty) sequence of $t_{\uparrow i'}$ transitions, for $i' \in W_1 \cup \dots \cup W_{j-1}$. Therefore, for all $i \in W_j \cup \dots \cup W_d$ we have $\sigma_j(i) = \sigma(i)$, and the same also holds for $j=1$ (since $\sigma_1 = \sigma$). In particular, $\sigma(W_j) = \sigma_j(W_j)$.

Now, since C_j only applies “push-to-the- $k(j-1) + 1$ st-index” transitions, and

$$\sigma_0(\{k(j-1) + 1\} \cup W_j \cup \dots \cup W_d) = \mathcal{R}_j \cup \dots \cup \mathcal{R}_d,$$

$$R \geq \begin{cases} \left(1 - \delta \lfloor \frac{1}{\delta} \rfloor\right) \log_2 \left(\left\lceil \frac{1}{\delta} \right\rceil + 1\right) + \delta \log_2 \left(\left(\left\lfloor \frac{1}{\delta} \right\rfloor + 1\right)!\right) \\ \quad + \left(1 - \delta \lfloor \frac{1}{\delta} \rfloor\right) \log_2 \left(\frac{\lfloor n/d \rfloor - 2}{\lfloor n/d \rfloor + 1}\right) - \delta & 4 < \lfloor 1/\delta \rfloor \equiv 0 \pmod{2}, \\ (1 - 4\delta) (\log_2(178)) + (5\delta - 1) \log_2(57) & \lfloor 1/\delta \rfloor = 4, \\ (1 - 2\delta) (3 - \log_2(3)) + \delta \log_2(3) & \lfloor 1/\delta \rfloor = 2, \\ \left(1 - \delta \lfloor \frac{1}{\delta} \rfloor\right) \log_2 \left(\left\lceil \frac{1}{\delta} \right\rceil + 1\right) + \delta \log_2 \left(\left(\left\lfloor \frac{1}{\delta} \right\rfloor + 1\right)!\right) \\ \quad + \left(\delta + \delta \lfloor \frac{1}{\delta} \rfloor - 1\right) \log_2 \left(\frac{\lfloor n/d \rfloor - 2}{\lfloor n/d \rfloor + 1}\right) - \delta & 5 < \lfloor 1/\delta \rfloor \equiv 1 \pmod{2}, \\ (1 - 5\delta) \log_2(315) + (6\delta - 1) \log_2(89) + 2 - 9\delta & \lfloor 1/\delta \rfloor = 5, \\ (1 - 3\delta) (\log_2(57) - 4) + 1 & \lfloor 1/\delta \rfloor = 3, \\ (1 - \delta) \log_2(3) & \lfloor 1/\delta \rfloor = 1. \end{cases} \quad (3)$$

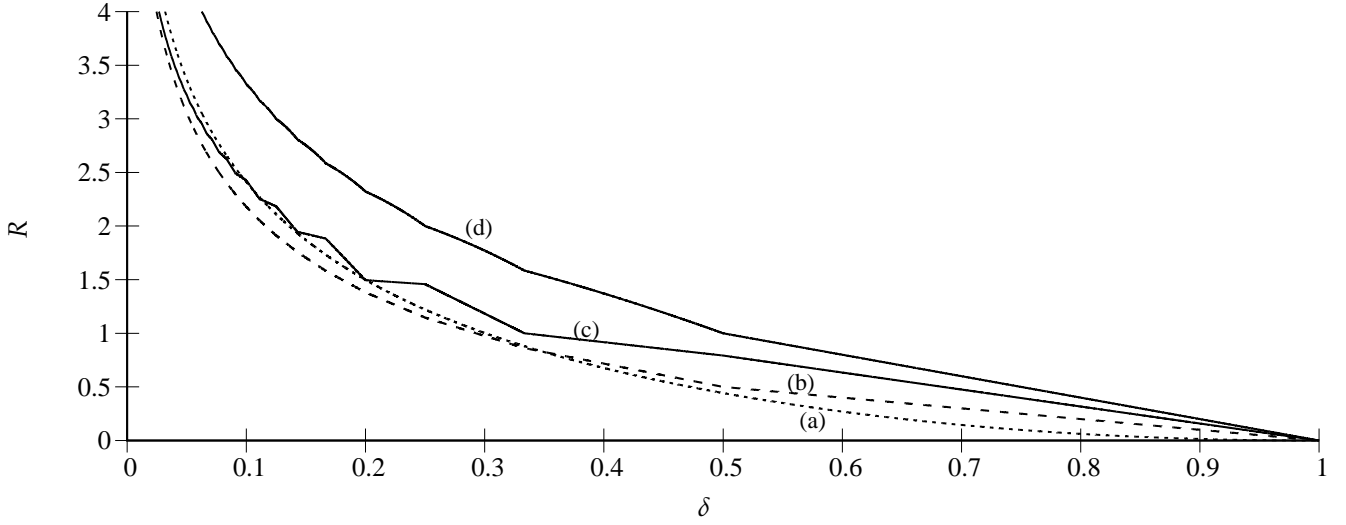


Figure 2. (a) The Gilbert-Varshamov-like lower bound of Lemma 22. (b) The rate of codes from Lemma 21 constructed in [32]. (c) A lower bound for the rate of codes C_1 from Construction A. (d) The upper bound of Lemma 20.

if for any $i \in W_j$ we have $\sigma(i) = \sigma_j(i) \notin \mathcal{R}_j$, then by necessity $\sigma(i) \in \mathcal{R}_{j'}$ for some $j' > j$. In particular, $\sigma(W_d) \subseteq \mathcal{R}_d$.

For all $j \in [d-1]$, we also consider σ_{j+1} , the C_{j+1} -parent of both σ and σ_j . Since C_{j+1} only applies “push-to-the- $kj+1$ st-index” transitions,

$$\begin{aligned} \sigma_{j+1}(\{k(j-1)+1\} \cup W_j \setminus \{kj+1\}) \\ = \sigma_0(\{k(j-1)+1\} \cup W_j \setminus \{kj+1\}) = \mathcal{R}_j. \end{aligned}$$

Finally, since σ_{j+1} is derived from σ_j by a sequence of $t_{k(j-1)+1 \uparrow i'}$ transitions for $i' \in W_j$, it follows that

$$\sigma_{j+1}(\{k(j-1)+1\} \cup W_j) = \sigma_j(\{k(j-1)+1\} \cup W_j)$$

thus

$$\begin{aligned} \sigma_j(W_j) \subseteq \sigma_{j+1}(\{k(j-1)+1\} \cup W_j) \\ = \mathcal{R}_j \cup \{\sigma_{j+1}(kj+1)\}. \end{aligned}$$

Noting that $|\sigma_j(W_j)| = |\mathcal{R}_j| = k$ and recalling that $\sigma(W_j) = \sigma_j(W_j)$, we are done. \blacksquare

Corollary 26 Let σ be a codeword of C_1 from Construction A, with $n = kd$. Then for each $j \in [d]$, there is a unique element $x_j^\sigma \in \mathcal{R}_j \cup \dots \cup \mathcal{R}_d$ satisfying

$$\sigma(W_j \cup \dots \cup W_d) = \mathcal{R}_j \cup \dots \cup \mathcal{R}_d \setminus \{x_j^\sigma\}.$$

Proof: The proposition follows from Lemma 25 for $j = d$ since $|\sigma(W_d)| = |W_d| = k-1 \leq |\mathcal{R}_d \cap \sigma(W_d)|$. Now suppose the proposition holds for $j+1$, and we prove that it holds for j .

We again observe by Lemma 25 that $|\mathcal{R}_j \cap \sigma(W_j)| \in \{k-1, k\}$. If $|\mathcal{R}_j \cap \sigma(W_j)| = k$, since $|\sigma(W_j)| = |W_j| = k$, then $\mathcal{R}_j = \sigma(W_j)$ and $x_j^\sigma = x_{j+1}^\sigma$ satisfies the claim.

Otherwise $\sigma(W_j) \setminus \mathcal{R}_j = \{y\}$ for some $y \in [n]$; it would suffice to show $y = x_{j+1}^\sigma$, since then $\mathcal{R}_j \setminus \sigma(W_j) = \{x_j^\sigma\}$ would satisfy the claim.

Consider then σ_j , the C_j -parent of σ . Note that $\sigma_j(W_j) = \sigma(W_j)$, and since C_j employs “push-to-the- $(k(j-1)+1)$ st-index” transitions only, and

$$\sigma_0(W_j \cup \dots \cup W_d) \subseteq \mathcal{R}_j \cup \dots \cup \mathcal{R}_d,$$

we know that $\sigma(W_j) \subseteq \mathcal{R}_j \cup \dots \cup \mathcal{R}_d$. We now use the induction hypothesis

$$\sigma(W_{j+1} \cup \dots \cup W_d) = (\mathcal{R}_{j+1} \cup \dots \cup \mathcal{R}_d) \setminus \{x_{j+1}^\sigma\},$$

and it follows that $\sigma(W_j) \subseteq \mathcal{R}_j \cup \{x_{j+1}^\sigma\}$, hence $y = x_{j+1}^\sigma$. ■

From now on, we denote $i_j^\sigma = \sigma^{-1}(x_j^\sigma)$. Another useful notation we shall employ is a function that quantizes any integer to the nearest integer leaving a residue of j modulo d . We denote this function by $q_d^j: \mathbb{Z} \rightarrow d\mathbb{Z} + j$, defined by

$$q_d^j(a) = \operatorname{argmin}_{b \in d\mathbb{Z} + j} |a - b|,$$

where we assume argmin returns a single value, and ties are broken arbitrarily.

For the decoding procedure description, let us fix the parameters $n = kd$, and the code C_1 from Construction A. Additionally, we denote by $\sigma \in C_1$ the transmitted permutation, by $\tau \in S_n$ the received permutation, and by $\hat{\sigma} \in S_n$ the decoded permutation. We denote the *decoding radius* by $t = \lfloor (d-1)/2 \rfloor$, and assume $d_\infty(\sigma, \tau) \leq t$.

We will decode τ iteratively by window, from W_1 to W_d . We shall make sure—inductively—that when we begin the process of decoding W_j , for some $j \in [d]$, we know i_j^σ . Initially, as mentioned, we set $j = 1$. Trivially, $i_1^\sigma = 1$.

Step I We set the decoding window

$$\hat{W}_j = W_j \cup \{i_j^\sigma\},$$

and naively decode \hat{W}_j by setting for all $i \in \hat{W}_j$,

$$\hat{\sigma}(i) = q_d^j(\tau(i)).$$

Lemma 27 After Step I, for all $i \in \hat{W}_j$ such that $\sigma(i) \in \mathcal{R}_j$ it holds that $\hat{\sigma}(i) = \sigma(i)$.

Proof: For all such i we have $\hat{\sigma}(i) \equiv \sigma(i) \pmod{d}$ and

$$\begin{aligned} |\hat{\sigma}(i) - \sigma(i)| &\leq |\hat{\sigma}(i) - \tau(i)| + |\tau(i) - \sigma(i)| \\ &= |q_d^j(\tau(i)) - \tau(i)| + |\tau(i) - \sigma(i)| \\ &\leq \lfloor d/2 \rfloor + t < d. \end{aligned}$$

Corollary 28 After Step I,

$$\hat{\sigma}(\hat{W}_j) = \hat{\sigma}(\hat{W}_j \setminus \{i_{j+1}^\sigma\}) = \mathcal{R}_j.$$

Proof: By Corollary 26 we know that $\mathcal{R}_j \subseteq \sigma(\hat{W}_j)$. We further recall that $\sigma(i_{j+1}^\sigma) \notin \mathcal{R}_j$, hence

$$\mathcal{R}_j \subseteq \sigma(\hat{W}_j \setminus \{i_{j+1}^\sigma\}),$$

and since

$$\left| \sigma(\hat{W}_j \setminus \{i_{j+1}^\sigma\}) \right| = \left| \hat{W}_j \setminus \{i_{j+1}^\sigma\} \right| = k = |\mathcal{R}_j|$$

we have equality. The claim now follows from Lemma 27. ■

Corollary 28 implies that after Step I, $\hat{\sigma}(\hat{W}_j)$ contains a unique element of \mathcal{R}_j which appears twice, and every other element appears exactly once; by Lemma 27 these other elements have been decoded correctly. Before we can continue inductively to decode W_{j+1} , it only remains to find i_{j+1}^σ ; the other instance in \hat{W}_j of $\hat{\sigma}(i_{j+1}^\sigma)$ we therefore also know to have been decoded correctly.

We shall identify i_{j+1}^σ using C^{aux} , the auxiliary $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code used in Construction A. By construction, if we examine σ_j , the C_j -parent of σ , then for all $i \in W_j$ we observe $\sigma(i) = \sigma_j(i)$, and $\sigma(i_j^\sigma) = \sigma_j(k(j-1) + 1)$. The ordering of the $k+1$ elements of $\sigma(\hat{W}_j) = \sigma_j(\{k(j-1) + 1\} \cup W_j)$ is then induced by a permutation of C^{aux} . We construct this induced permutation from the auxiliary code C^{aux} , which we denote $\hat{\pi} \in S_{k+1}$. We first define a simple bijection $\alpha_j: \mathcal{R}_j \rightarrow [k]$, which is the inverse of the enumeration of \mathcal{R}_j given by the arbitrary initial order of elements in σ_0 used in Construction A, e.g., in the simple case $n = kd$,

$$\alpha_j(m) = \begin{cases} \lfloor \frac{m}{d} \rfloor & j < m \in \mathcal{R}_j, \\ k & m = j. \end{cases}$$

With α_j we define $\hat{\pi}$ as,

$$\hat{\pi}(i) = \begin{cases} \alpha_j(\hat{\sigma}(i_j^\sigma)) & i = 1; \\ \alpha_j(\hat{\sigma}(k(j-1) + i)) & i \in \{2, 3, \dots, k+1\}, \end{cases}$$

and note that—as it currently stands— $\hat{\pi}$ is not a permutation of $[k+1]$ because its range is $[k]$ and some unique $a \in [k]$ has two distinct pre-images.

Theorem 29 Let $s, t \in [k+1]$ be the unique pair of indices such that $\hat{\pi}(s) = \hat{\pi}(t) = a \in [k]$. There is a unique way to redefine $\hat{\pi} \upharpoonright_{\{s,t\}}$ (the restriction of $\hat{\pi}$ to $\{s, t\}$) as a bijection onto $\{a, k+1\}$ that yields $\hat{\pi} \in C^{\text{aux}}$. Furthermore, if we define $I_j: [k+1] \times [n] \rightarrow [n]$ by

$$I_j(q, r) = \begin{cases} r & q = 1, \\ k(j-1) + q & \text{otherwise} \end{cases}$$

then after performing that correction

$$i_{j+1}^\sigma = I_j(\hat{\pi}^{-1}(k+1), i_j^\sigma).$$

■

Proof: First, arbitrarily set $\hat{\pi}(t) = k+1$, where $t > s$. Once corrected, $\hat{\pi} \in S_{k+1}$ by Corollary 28 and because $\alpha_j: \mathcal{R}_j \rightarrow [k]$ is a bijection.

Now, we take $\pi \in C^{\text{aux}}$ which generates σ_j in the recursion step of Construction A—while constructing C_j —from its C_{j+1} -parent. Hence

$$\pi(i) = \begin{cases} \alpha_j(x_j^\sigma) & i = 1, \\ \alpha_j(\sigma(k(j-1) + i)) & i \in \{2, 3, \dots, k+1\}, \end{cases}$$

and therefore either $\hat{\pi} = \pi$ or $\hat{\pi} = (k+1, a) \circ \pi$. Crucially, we observe that in the latter case $\hat{\pi} \notin C^{\text{aux}}$ since C^{aux} is a $G_{\uparrow}^{\text{aux}}(k+1, \tilde{M}_{k+1})$ code and $\pi \in C^{\text{aux}}$; we utilize `ValidAux` to discover whether our original arbitrary correction should be reversed.

To complete the proof, we note by the recursion step of Construction A that, indeed, $i_{j+1}^\sigma = I_j(\tau^{-1}(k+1), i_j^\sigma)$. ■

We can therefore complete our iterative decoding round with the following step.

Step II We construct $\hat{\pi}$ as described, identify $s, t, s < t$, and arbitrarily correct $\hat{\pi}(t) = k+1$. We test $\text{ValidAux}(\hat{\pi})$: if true, we have $i_{j+1}^\sigma = I_j(t, i_j^\sigma)$; otherwise, it holds that $i_{j+1}^\sigma = I_j(s, i_j^\sigma)$.

Finally, observe that when decoding W_d it's known that $\sigma(\hat{W}_d) = \mathcal{R}_d$, hence by Lemma 27 \hat{W}_d is decoded correctly, and we need not (and indeed cannot) perform Step II.

Example 30 We shall demonstrate the decoding process assuming once again $n = kd$ for simplicity, and using the parameters $d = 3$ (hence $t = 1$), $k = 2$ and code constructed in Example 18. Recall that the $G_{\uparrow}^{\text{aux}}(3, 3)$ code used in that example is

$$C^{\text{aux}} = \{[1, 2, 3], [3, 1, 2], [2, 3, 1]\}.$$

We choose the transmitted codeword $\sigma = [1, 2, 4, 6, 5, 3]$, and a noisy received permutation $\tau = [1, 3, 4, 5, 6, 2]$.

We start by defining $i_1 = 1$ and observing (by abuse of the vector notation) $\tau \upharpoonright_{\hat{W}_1} = [1; 3, 4]$ (the first element is differentiated because—generally although never when $j = 1$ —it does not immediately precede the rest in τ 's vector notation).

Since $j = 1$, we define $\hat{\sigma} \upharpoonright_{\hat{W}_1} = [1; 4, 4]$. This leads us to construct $\hat{\pi} = [2, 1, 3] \notin C^{\text{aux}}$, so we instead correct $\hat{\pi} = [2, 3, 1]$ and define $i_2 = 2$. (So far we have $\hat{\sigma} = [1, \underline{4}, 4, \cdot, \cdot, \cdot]$, where an underline marks i_{j+1}^σ .)

Next, we have $\tau \upharpoonright_{\hat{W}_2} = [3; 5, 6]$, which ($j = 2$) we decode $\hat{\sigma} \upharpoonright_{\hat{W}_2} = [2; 5, 5]$. This again generates $\hat{\pi} = [2, 1, 3] \notin C^{\text{aux}}$, and we correct in similar fashion to $\hat{\pi} = [2, 3, 1]$ and define $i_2 = 4$. (Up to this point, we have $\hat{\sigma} = [1, \underline{4}, 2, 4, \underline{5}, 5, \cdot]$.)

Finally, we have $\tau \upharpoonright_{\hat{W}_3} = [5; 2]$ and since $j = 3$ we decode $\hat{\sigma} \upharpoonright_{\hat{W}_3} = [6; 3]$, and overall $\hat{\sigma} = [1, 2, 4, \underline{5}, 6, 5, 3] = \sigma$. □

Example 31 We present another example, intended to demonstrate the process in more detail, for which we depart from the parameters used in Example 18 by setting $d = 5$ (allowing for $t = 2 \leq \lfloor (d-1)/2 \rfloor$), $k = 3$. In each recursion step of Construction A the $G_{\uparrow}^{\text{aux}}(4, 8)$ code used is C^{aux} presented in Example 9.

The codeword

$$\sigma = [11, 1, 8, 6, 7, 2, 12, 13, 3, 5, 9, 14, 4, 10, 15]$$

appears in the code generated in this case, as can be seen by identifying its C_5 , C_4 , C_3 , and C_2 parents as, respectively,

$$\sigma_5 = [6, 11, 1, 7, 12, 2, 8, 13, 3, 9, 14, 4, 5, 10, 15],$$

$$\sigma_4 = [6, 11, 1, 7, 12, 2, 8, 13, 3, 5, 9, 14, 4, 10, 15],$$

$$\sigma_3 = \sigma_4,$$

$$\sigma_2 = [6, 11, 1, 8, 7, 2, 12, 13, 3, 5, 9, 14, 4, 10, 15].$$

We choose

$$\tau = [12, 3, 9, 7, 5, 2, 11, 15, 1, 6, 8, 13, 4, 10, 14]$$

to be the noisy version of the transmitted codeword σ , and verify that $d_\infty(\tau, \sigma) = 2 = t$.

Beginning with $j = 1$, we have $\tau \upharpoonright_{\hat{W}_1} = [12; 3, 9, 7]$, which we decode $\hat{\sigma} \upharpoonright_{\hat{W}_1} = [11; 1, 11, 6]$, generating $\hat{\pi} = [2, 3, 2, 1]$ which is corrected to $\hat{\pi} = [2, 3, 4, 1] \in C^{\text{aux}}$. We identify $i_2 = 3$, and keep

$$\hat{\sigma} = [11, 1, \underline{11}, 6, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot].$$

Next, for $j = 2$, observe that $\tau \upharpoonright_{\hat{W}_2} = [9; 5, 2, 11]$, and we decode $\hat{\sigma} \upharpoonright_{\hat{W}_2} = [7; 7, 2, 12]$. This generates $\hat{\pi} = [1, 1, 3, 2]$, which we initially correct to $\hat{\pi} = [1, 4, 3, 2] \notin C^{\text{aux}}$, so (skip correcting $\hat{\pi}$, as it has no further consequence) $i_3 = i_2 = 3$ instead of $i_3 = 4$. We summarize

$$\hat{\sigma} = [11, 1, \underline{11}, 6, 7, 2, 12, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot, \cdot].$$

We turn to \hat{W}_3 and see that $\tau \upharpoonright_{\hat{W}_3} = [9; 15, 1, 6]$, decoded to $\hat{\sigma} \upharpoonright_{\hat{W}_3} = [8; 13, 3, 8]$. We generate $\hat{\pi} = [1, 2, 3, 1]$ and correct it to $\hat{\pi} = [1, 2, 3, 4] \in C^{\text{aux}}$, indicating that $i_4 = 10$. We now have

$$\hat{\sigma} = [11, 1, \underline{7}, 8, 6, 7, 2, 12, 13, 3, \underline{8}, \cdot, \cdot, \cdot, \cdot].$$

Moving on to $j = 4$, while decoding \hat{W}_4 we note $\tau \upharpoonright_{\hat{W}_4} = [6; 8, 13, 4]$, which we decode as $\hat{\sigma} \upharpoonright_{\hat{W}_4} = [4; 9, 14, 4]$. This generates $\hat{\pi} = [3, 1, 2, 3]$ which is corrected to $\hat{\pi} = [3, 1, 2, 4] \notin C^{\text{aux}}$. We therefore define $i_5 = i_4 = 10$ instead of $i_5 = 13$. Up to now,

$$\hat{\sigma} = [11, 1, 8, 6, 7, 2, 12, 13, 3, \underline{8}, 9, 14, 4, \cdot, \cdot].$$

Finally, $j = 5$, and we get $\tau \upharpoonright_{\hat{W}_5} = [4; 10, 14]$ which we decode to $\hat{\sigma} \upharpoonright_{\hat{W}_5} = [5; 10, 15]$, and overall

$$\hat{\sigma} = [11, 1, 8, 6, 7, 2, 12, 13, 3, \underline{8}, 5, 9, 14, 4, 10, 15] = \sigma. \quad \square$$

The decoding algorithm is formalized in $\text{Decode}(\tau)$. With appropriate simple data structures, the algorithm requires $O(kd) = O(n)$ steps. We assume simple integer operations to take constant-time.

VI. RANKING AND UNRANKING

In this section we discuss the process of encoding data $m \in \{0, 1, \dots, |C_1| - 1\}$ to a codeword $\sigma \in C_1$, which is also known as *unranking* m , and the inverse process of *ranking* $\sigma \in C_1$, i.e., obtaining its rank in the code. Throughout this section, C_1 stands for the code obtained via Construction A.

Due to the nature of our construction, performing these tasks with the codes generated by Theorem 17 is reliant on our ability to do the same with the codes provided by Lemma 7 and Corollary 14. We therefore recall the following known result.

Lemma 32 [21] The complete $G_{\uparrow}(n, n!)$ codes provided by Lemma 7 has a ranking algorithm operating in $O(n)$ steps, and an unranking scheme operating in $O(n^2)$ steps.

This gives rise to the following corollary.

```

Function Decode ( $\tau$ )
  input :  $\tau \in S_{kd}$  satisfying  $d_\infty(\tau, C_1) \leq t \leq \lfloor (d-1)/2 \rfloor$ .
  output :  $\hat{\sigma} \in C_1$  such that  $d_\infty(\tau, \hat{\sigma}) \leq t$ .
  1  $i \leftarrow 1$ 
  2 for  $j = 1, 2, \dots, d-1$  do
    /* Naively decode  $\hat{W}_j$  */
  3  $\hat{\sigma}(i) \leftarrow q_d^j(\tau(i))$ 
  4  $\hat{\pi}(1) \leftarrow \alpha_j(\hat{\sigma}(i))$ 
  5 for  $r = 2, \dots, k+1$  do
  6    $m \leftarrow q_d^j(\tau(k(j-1) + r))$ 
  7   if  $\hat{\sigma}^{-1}(m)$  is already set then
  8      $\hat{\pi}(r) \leftarrow k+1$ 
  9      $a \leftarrow \alpha_j(m)$ 
 10   else
 11      $\hat{\pi}(r) \leftarrow \alpha_j(m)$ 
 12    $\hat{\sigma}(r) \leftarrow m$ 
  /* Define  $i_{j+1}$  */
 13 if ValidAux( $\hat{\pi}$ ) then
 14    $i \leftarrow I(\hat{\pi}^{-1}(k+1), i)$ 
 15 else
 16    $i \leftarrow I(\hat{\pi}^{-1}(a), i)$ 
  /* Decode  $\hat{W}_d$  */
 17  $\hat{\sigma}(i) \leftarrow q_d^d(\tau(i))$ 
 18 for  $r = 2, \dots, k$  do
 19    $\hat{\sigma}(r) \leftarrow q_d^j(\tau(k(d-1) + r))$ 
 20 return  $\hat{\sigma}$ 

```

Corollary 33 The $G_{\uparrow}^{\text{aux}}(2m, \frac{|S_{2m}|}{2^{m-1}})$ codes generated by Theorem 8 can be ranked in $O(m)$ operations and unranked in $O(m^2)$ operations.

Proof: Ranking a permutation σ in the code may proceed by finding the cyclic shift required for $[2m, 1]$ to be the first two elements. After removing these two first elements, and then reversing the permutation we may use a ranking algorithm from Lemma 32. A simple combination of the results gives the required ranking of σ . By Lemma 32, the entire procedure takes $O(m)$ operations. A symmetric argument gives an $O(m^2)$ algorithm for unranking. ■

Unfortunately, no ranking and unranking schemes are known for parity-preserving $G_{\uparrow}(2m+1, M_{2m+1})$ codes provided by Lemma 6 (developed in [18]), or previous constructions presented in [20], [38]. Consequentially, we rely on Theorem 8 instead of Theorem 13 for even sized congruence classes. In the case of odd sized classes, we can leverage the following codes.

Lemma 34 [36] For all $m \geq 1$ there exist parity-preserving $G_{\uparrow}(2m+1, \hat{M}_{2m+1})$ codes with sizes

$$\hat{M}_{2m+1} = \left(\frac{(2m)!}{m!} \right)^2 \frac{(2m+1)}{2^{2m}} = \frac{(2m)!}{m! 2^{2m}} |S_{2m+1}|.$$

These codes can be ranked and unranked in $O(m^2)$ operations.

We summarize those observations in the following corollary.

Corollary 35 For all $k \geq 3$ there exist a $G_{\uparrow}^{\text{aux}}(k, \hat{M}_k)$ code starting with Id and a $t_{\uparrow k}$ transition, which have ranking and unranking schemes operating in $O(k^2)$ steps, where

$$\hat{M}_k = \begin{cases} \frac{(k-1)!}{((k-1)/2)! 2^{k-1}} k! & k \equiv 1 \pmod{2}; \\ \frac{k!}{(k-1)!} & k \equiv 0 \pmod{2}. \end{cases}$$

Note that we can now replace Corollary 14 by Corollary 35 in Construction A to obtain codes which we shall denote \hat{C}_1 , and each auxiliary code on a congruence class of size $k > 1$ contributes to $|\hat{C}_1|$ a multiplicative factor of

$$\hat{M}_{k+1} = \begin{cases} \frac{k!}{((k/2)! 2^{k/2})} (k+1)! & k \equiv 0 \pmod{2}, \\ \frac{(k+1)!}{k} & k \equiv 1 \pmod{2}. \end{cases}$$

We also note, using Stirling's approximation

$$e^{\frac{1}{12n+1}} < \frac{n! e^n}{n^n \sqrt{2\pi n}} < e^{\frac{1}{12n}},$$

that

$$\frac{k!}{(k/2)! 2^{k/2}} > \sqrt{\frac{2}{\pi k}} \left(e^{1 - \frac{1}{4 + \frac{1}{3k}}} \right)^{-\frac{1}{3k}} > \sqrt{\frac{2}{\pi k}} e^{-1/(4k)}.$$

We can then recalculate code size, in the case of $\lfloor \frac{n}{d} \rfloor \equiv 0 \pmod{2}$:

$$\begin{aligned} |\hat{C}_1| &= \left(\frac{(\lceil n/d \rceil + 1)!}{\lfloor n/d \rfloor!} \right)^{n \bmod d} \cdot \lfloor \frac{n}{d} \rfloor! \\ &\quad \cdot \left(\frac{\lfloor n/d \rfloor!}{((\lfloor n/d \rfloor)/2)! 2^{\lfloor n/d \rfloor}} \left(\lfloor \frac{n}{d} \rfloor + 1 \right)! \right)^{d - (n \bmod d) - 1} \\ &> \lfloor \frac{n}{d} \rfloor!^{n \bmod d} \lfloor \frac{n}{d} \rfloor!^{d - (n \bmod d)} \\ &\quad \cdot \left(1 + \frac{1}{\lfloor n/d \rfloor} \right)^{n \bmod d} \cdot \left(\lfloor \frac{n}{d} \rfloor + 1 \right)^{(d-1) - (n \bmod d)} \\ &\quad \cdot \left(\sqrt{\frac{2}{\pi \lfloor n/d \rfloor}} e^{-1/(4\lfloor n/d \rfloor)} \right)^{(d-1) - (n \bmod d)}, \end{aligned}$$

and when $\lfloor \frac{n}{d} \rfloor \equiv 1 \pmod{2}$:

$$\begin{aligned} |\hat{C}_1| &= \left(\frac{\lceil n/d \rceil!}{((\lceil n/d \rceil)/2)! 2^{\lceil n/d \rceil}} \left(\lceil \frac{n}{d} \rceil + 1 \right)! \right)^{n \bmod d} \\ &\quad \cdot \lfloor \frac{n}{d} \rfloor! \cdot \left(\frac{(\lfloor n/d \rfloor + 1)!}{\lfloor n/d \rfloor!} \right)^{d - (n \bmod d) - 1} \\ &> \lfloor \frac{n}{d} \rfloor!^{n \bmod d} \lfloor \frac{n}{d} \rfloor!^{d - (n \bmod d)} \\ &\quad \cdot \left(\lceil \frac{n}{d} \rceil + 1 \right)^{n \bmod d} \cdot \left(1 + \frac{1}{\lfloor n/d \rfloor} \right)^{(d-1) - (n \bmod d)} \\ &\quad \cdot \left(\sqrt{\frac{2}{\pi \lfloor n/d \rfloor}} e^{-1/(4\lfloor n/d \rfloor)} \right)^{n \bmod d}, \end{aligned}$$

and we note that in the special case $\lfloor \frac{n}{d} \rfloor = 1$ we have $|\hat{C}_1| = |C_1|$.

Function Rank(σ)
<pre> input : $\sigma \in \hat{C}_1$. output : $m \in \{0, 1, \dots, \hat{C}_1 - 1\}$ which is the rank of σ in \hat{C}_1. /* Build a permutation $\pi_d \in S_k$ */ 1 for $i \in W_d$ do 2 $\pi_d[i - k(d - 1)] \leftarrow \alpha_d(\sigma[i])$ 3 $\pi_d[1] \leftarrow [k] \setminus \pi_d[2, \dots, k]$ 4 $m \leftarrow ((\text{RankComplete}(\pi_d) - 1) \bmod k!)$ 5 for $j = d - 1, \dots, 1$ do 6 /* Build a permutation $\pi_j \in S_{k+1}$ */ 7 for $i \in W_j$ do 8 $\pi_j[i - k(j - 1)] \leftarrow \alpha_j(\sigma[i])$ 9 $\pi_j[1] \leftarrow [k + 1] \setminus \pi_j[2, \dots, k + 1]$ 10 $m \leftarrow m \cdot \hat{M}_{k+1} + ((\text{RankAux}(\pi_j) - 1) \bmod \hat{M}_{k+1})$ 11 return $(m + 1) \bmod \hat{C}_1$ </pre>

We likewise observe the rates of codes based on Corollary 35, and find for $\lfloor 1/\delta \rfloor \equiv 0 \pmod{2}$

$$\begin{aligned}
\hat{R} &\geq \left(1 - \delta \left\lfloor \frac{1}{\delta} \right\rfloor\right) \log_2 \left(\left\lfloor \frac{1}{\delta} \right\rfloor! \left(1 + \frac{1}{\lfloor 1/\delta \rfloor}\right) \right) \\
&\quad + \left(\delta + \delta \left\lfloor \frac{1}{\delta} \right\rfloor - 1\right) \log_2 \left(\left(\left\lfloor \frac{1}{\delta} \right\rfloor + 1 \right)! \right) \\
&\quad - \frac{1}{2} \left(\delta + \delta \left\lfloor \frac{1}{\delta} \right\rfloor - 1\right) \\
&\quad \cdot \left(\log_2 \left\lfloor \frac{1}{\delta} \right\rfloor + \frac{\log_2(e)}{2 \lfloor 1/\delta \rfloor} + \log_2(\pi) - 1 \right) - o(1),
\end{aligned}$$

and for $\lfloor 1/\delta \rfloor \equiv 1 \pmod{2}$

$$\begin{aligned}
\hat{R} &\geq \left(1 - \delta \left\lfloor \frac{1}{\delta} \right\rfloor\right) \log_2 \left(\left(\left\lfloor \frac{1}{\delta} \right\rfloor + 1 \right)! \right) \\
&\quad + \left(\delta + \delta \left\lfloor \frac{1}{\delta} \right\rfloor - 1\right) \log_2 \left(\left\lfloor \frac{1}{\delta} \right\rfloor! \left(1 + \frac{1}{\lfloor 1/\delta \rfloor}\right) \right) \\
&\quad - \frac{1}{2} \left(1 - \delta \left\lfloor \frac{1}{\delta} \right\rfloor\right) \\
&\quad \cdot \left(\log_2 \left\lfloor \frac{1}{\delta} \right\rfloor + \frac{\log_2(e)}{2 \lfloor 1/\delta \rfloor} + \log_2(\pi) - 1 \right) - o(1).
\end{aligned}$$

The losses in asymptotic rate are shown in Figure 3. We observe in particular that we still manage to achieve better rates than previously known error-correcting codes (without the Gray property), even with the significantly smaller $G_{\uparrow}^{\text{aux}}(k, \hat{M}_k)$ of Corollary 35.

Let us denote by $\text{RankComplete}(\pi)$, $\text{UnrankComplete}(m)$ the ranking and unranking procedures for the complete codes from Lemma 32. Additionally, let $\text{RankAux}(\pi)$ and $\text{UnrankAux}(m)$ denote the ranking and unranking procedures for the auxiliary codes of Corollary 35. We can readily take advantage of \hat{C}_1 's tiered structure to use these functions in order to perform the same tasks for our construction. We include pseudo-code for these algorithms, which we call $\text{Rank}(\sigma)$ and $\text{Unrank}(m)$, for completeness. As before, we assume $n = kd$ to simplify the presentation.

Theorem 36 For the code \hat{C}_1 of length $n = kd$, the algorithms $\text{Rank}(\sigma)$, $\text{Unrank}(m)$ operate in $O(k^2d)$ steps.

Proof: Both algorithms perform a single loop over all indices of σ , making simple integer operations, which

Function Unrank(m)
<pre> input : $m \in \{0, 1, \dots, \hat{C}_1 - 1\}$. output : $\sigma \in \hat{C}_1$ with rank m in \hat{C}_1. /* Convert m to local ranks $R[1, 2, \dots, d]$ */ 1 $m \leftarrow ((m - 1) \bmod \hat{C}_1)$ 2 for $i = 1, 2, \dots, d - 1$ do 3 $R[i] \leftarrow ((m + 1) \bmod \hat{M}_{k+1})$ 4 $m \leftarrow \lfloor m / \hat{M}_{k+1} \rfloor$ 5 $R[d] \leftarrow ((m + 1) \bmod k!)$ 6 /* Construct σ */ 7 $\pi_d \leftarrow \text{UnrankComplete}(R[d])$ 8 for $i \in W_d$ do 9 $\sigma[i] \leftarrow \pi_d[i - k(d - 1)] \cdot d$ 10 $x \leftarrow \pi_d[1] \cdot d$ 11 for $j = d - 1, \dots, 1$ do 12 $\pi_j \leftarrow \text{UnrankAux}(R[j])$ 13 for $i \in W_j$ do 14 if $\pi_j[i] = k + 1$ then 15 $\sigma[i] \leftarrow x$ 16 else 17 $\sigma[i] \leftarrow \pi_j[i - k(j - 1)] \cdot d + j$ 18 if $\pi_j[1] \neq k + 1$ then 19 $x \leftarrow \pi_j[1] \cdot d + j$ 20 $\sigma[1] \leftarrow x$ 21 return σ </pre>

requires $O(n)$ steps. They also make a call to one of $\text{RankComplete}(\pi)$, $\text{UnrankComplete}(m)$ and $(d - 1)$ calls to one of $\text{RankAux}(\pi)$, $\text{UnrankAux}(m)$, costing $O(k^2)$ operations each. ■

We also note in particular that in the regime $d = \Theta(n)$, we have $k = \Theta(1)$, and Theorem 36 yields linear run-time $O(n)$.

VII. SNAKE-IN-THE-BOX CODES IN S_{2m+2}

As mentioned before in Section III, the issue of asymmetry between “push-to-the-top” codes in the symmetric group of odd and even orders has also frustrated research into error-detecting codes under the Kendall τ -metric in the past.

The Kendall τ -metric [23] on S_n is defined as

$$d_{\mathcal{K}}(\sigma, \tau) = |\{(i, j) \mid \sigma(i) < \sigma(j) \wedge \tau(i) > \tau(j)\}|.$$

Informally, as noted in [22], it measures the minimal number of adjacent transpositions required to transform one permutation into the other, that is, the minimal r such that

$$\sigma = \tau \circ (i_1, i_1 + 1) \circ (i_2, i_2 + 1) \circ \dots \circ (i_r, i_r + 1)$$

for some $i_1, i_2, \dots, i_r \in [n - 1]$. An (n, M, \mathcal{K}) -snake, or \mathcal{K} -snake for short, is a single-error-detecting rank-modulation Gray code of size M , or more formally, a $G_{\uparrow}(n, M)$ code \mathcal{C} such that for all $\sigma, \tau \in \mathcal{C}$, $\sigma \neq \tau$, it holds that $d_{\mathcal{K}}(\sigma, \tau) \geq 2$. Put differently, for no $i \in [n - 1]$ does it hold that $\sigma = \tau \circ (i, i + 1)$.

The authors have shown in [36][Thm. 17] that any \mathcal{K} -snake $\mathcal{C} \subseteq S_n$ which employs a “push-to-the-top” transition on an even index $t_{\uparrow 2m}$ —for any $m \in \lfloor \frac{n}{2} \rfloor$ —must satisfy $|\mathcal{C}| \leq \frac{n!}{2} - \Theta(n)$. Horovitz and Etzion posited in [20] that \mathcal{K} -snakes in S_{2m+2} do not exceed the size of those in S_{2m+1} , a conjecture refuted when Zhang and Ge demonstrated in [37] the existence of \mathcal{K} -snakes in S_{2m+2} of size $\frac{(2m+2)!}{4}$.

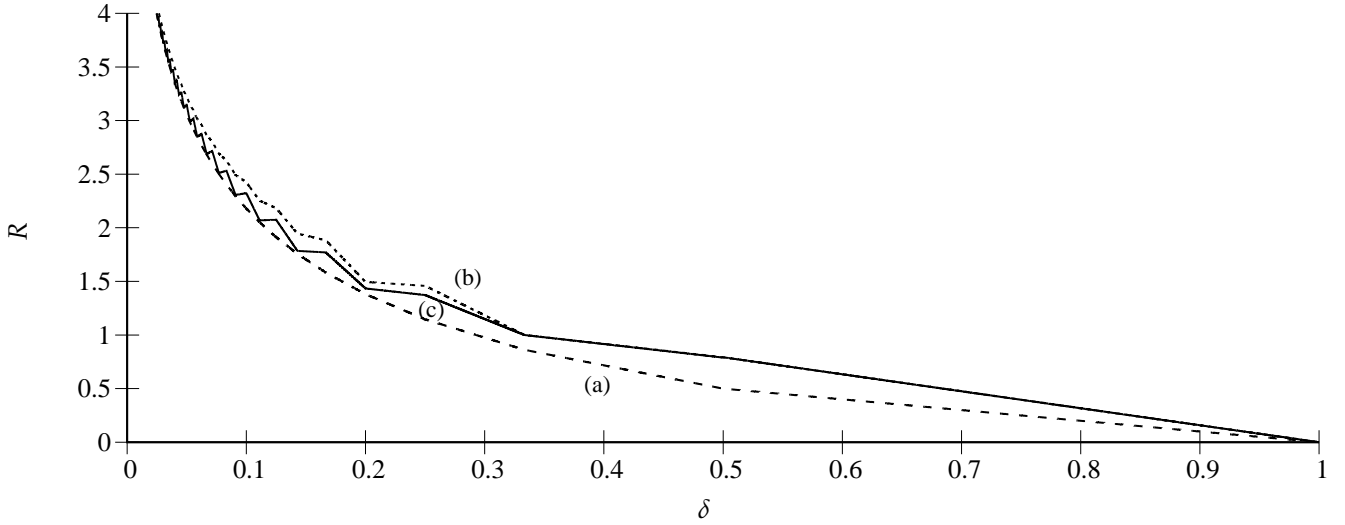


Figure 3. (a) The rate of codes from Lemma 21 constructed in [32]. (b) The rate of codes C_1 from Construction A. (c) The rate of codes \hat{C}_1 constructed using auxiliary codes from Corollary 35.

Concurrently and independently, Holroyd conjectured in [18] that \mathcal{K} -snakes can be found in S_{2m+2} with size greater than $\frac{(2m+2)!}{2} - O(m^2)$.

A resemblance is evident in the definitions of (n, M, \mathcal{K}) -snakes and $G_{\uparrow}^{\text{aux}}(n, M)$ codes, which is reinforced by the observations that, similarly to properties seen in Section III, any parity-preserving $G_{\uparrow}(n, M)$ code is an (n, M, \mathcal{K}) -snake (see [36][Lem. 5]), and any (n, M, \mathcal{K}) -snake satisfies $M \leq \frac{n!}{2}$ (see [36][Thm. 15]).

We wish to demonstrate how the principles behind Theorem 13 can be applied to the construction of a \mathcal{K} -snake in S_{2m+2} of size $M \approx \frac{(2m+2)!}{2}$.

Lemma 37 [20, Thm. 18] [38] For $m \geq 2$, there exist parity-preserving $G_{\uparrow}(2m+1, M_{2m+1})$ codes with

$$M_{2m+1} = |A_{2m+1}| - (2m-1) = \frac{(2m+1)!}{2} - (2m-1).$$

In particular, such a code C was constructed such that, as a group,

$$C = A_{2m+1} \setminus \{t_{\uparrow 2m-1}^q \sigma\}_{q=0}^{2m-2}$$

for some $\sigma \in A_{2m+1}$. Finally, C only employed $t_{\uparrow 2m-1}$, $t_{\uparrow 2m+1}$.

As before, we fix $m \geq 2$. We also reuse

$$\begin{aligned} \varphi(\pi) &= t_{\uparrow 2m+2}^2 \circ t_{\uparrow 2m-1}^{-1}(\pi) \\ &= \pi \circ (1, 2m+1)(2m+2, 2m, 2m-1, \dots, 2) \end{aligned}$$

and the permutations $\hat{\pi}_r = \varphi^r(\text{Id})$.

Theorem 38 For all $r \geq 0$ a parity-preserving $G_{\uparrow}(2m+2, \frac{(2m+1)!}{2} - (2m-1))$ code \hat{P}_r exists which satisfy:

- 1) The first permutation in \hat{P}_r is $\hat{\pi}_r$.
- 2) The last permutation in \hat{P}_r is $t_{\uparrow 2m-1}^{-1} \hat{\pi}_r$.

- 3) For all $\pi \in \hat{P}_r$ it holds that

$$\begin{aligned} \pi(2m+2) &= \hat{\pi}_r(2m+2) \\ &= \begin{cases} 2m+2 & r \equiv 0 \pmod{2m}, \\ 2m+1 - (r \bmod 2m) & r \not\equiv 0 \pmod{2m}. \end{cases} \end{aligned}$$

- 4) $\tilde{\sigma}_r \notin \hat{P}_r$, where we denote

$$\tilde{\sigma}_r = (t_{\uparrow 2m+2}^{-1} \hat{\pi}_r) \circ (2m+1, 2m+2)$$

(and observe $\tilde{\sigma}_r = t_{\uparrow 2m+1}^{-1}(\hat{\pi}_r)$, hence in particular $\tilde{\sigma}_r(2m+2) = \hat{\pi}_r(2m+2)$).

Proof: By Lemma 37 we know that there exist a parity-preserving $G_{\uparrow}(2m+1, M_{2m+1})$ code P such that, as a group,

$$P = S_{2m+1} \setminus \{t_{\uparrow 2m-1}^q \sigma\}_{q=0}^{2m-2}$$

for some $\sigma \in A_{2m+1}$. We also know that P only employs $t_{\uparrow 2m-1}$, $t_{\uparrow 2m+1}$ transitions.

We apply its generating sequence to $\hat{\pi}_r$ to generate the $G_{\uparrow}(2m+2, M_{2m+1})$ code \hat{P} , which employs only $t_{\uparrow 2m-1}$, $t_{\uparrow 2m+1}$ transitions (in particular, it never employs $t_{\uparrow 2m+2}$, hence point 3 is established), and note that as a group

$$\begin{aligned} \hat{P} &= \{\tau \in A_{2m+2} \mid \tau(2m+2) = \hat{\pi}_r(2m+2)\} \\ &\quad \setminus \{t_{\uparrow 2m-1}^q \hat{\sigma}\}_{q=0}^{2m-2} \end{aligned}$$

for some $\hat{\sigma} \in A_{2m+2}$, satisfying $\hat{\sigma}(2m+2) = \hat{\pi}_r(2m+2)$.

Denote $\hat{P} = (c_j)_{j=1}^{M_{2m+1}}$. We modify our code by defining

$$\hat{P}_r = (c'_j)_{j=1}^{M_{2m+1}} = (\tilde{\sigma}_r \hat{\sigma}^{-1} c_j)_{j=1}^{M_{2m+1}},$$

which is still a $G_{\uparrow}(2m+2, M_{2m+1})$ since “push-to-the-top” transitions are group-actions by right-multiplication. Moreover, since $\tilde{\sigma}_r(2m+2) = \hat{\sigma}(2m+2) = \hat{\pi}_r(2m+2)$, as a group we have

$$\begin{aligned} \hat{P}_r &= \{\tau \in A_{2m+2} \mid \tau(2m+2) = \hat{\pi}_r(2m+2)\} \\ &\quad \setminus \{t_{\uparrow 2m-1}^q \tilde{\sigma}_r\}_{q=0}^{2m-2}. \end{aligned}$$

Note in particular that

$$\tilde{\sigma}_r(2m+1) = \hat{\pi}_r(1) \neq \hat{\pi}_r(2m+1),$$

hence $\hat{\pi}_r \in \hat{P}_r$. In addition, point 4 is thus substantiated.

Finally, $t_{\uparrow 2m+1}^{-1}(\hat{\pi}_r) = \tilde{\sigma}_r \notin \hat{P}_r$ implies that $\hat{\pi}_r$ must necessarily be preceded in \hat{P}_r by $t_{\uparrow 2m-1}$, which substantiates point 2 (after a proper cyclic shift of \hat{P}_r). ■

As in Section III, $\hat{P}_r \subseteq A_{2m+2}$ for all r . We construct a $(2m+2, M, \mathcal{K})$ -snake by stitching together $\hat{P}_0, \hat{P}_1, \dots, \hat{P}_{2m-1}$ in the following lemma.

Lemma 39 For all $r \geq 0$, we may concatenate \hat{P}_r, \hat{P}_{r+1} into a (non-cyclic) “push-to-the-top” code by applying the transitions $t_{\uparrow 2m+2}, t_{\uparrow 2m+2}$ to the last permutation of \hat{P}_r , which is $t_{\uparrow 2m-1}^{-1}\hat{\pi}_r$.

The only odd permutation in the resulting code is then

$$\beta_{r+1} = t_{\uparrow 2m+2}^{-1}(\hat{\pi}_{r+1}),$$

which we again call the $(r+1)$ -bridge.

Proof: Exactly as in the proof of Lemma 11, given that P_r, \hat{P}_r are parity-preserving, and have the same first and last permutations. ■

Again, similarly to Section III, Lemma 39 can be used iteratively to cyclically concatenate $\hat{P}_0, \hat{P}_1, \dots, \hat{P}_{2m-1}$, with a single odd permutation—the r -bridge—between $\hat{P}_{(r-1) \bmod 2m}, \hat{P}_r$. Let us prove that fact in the following theorem.

Theorem 40 There exists a $(2m+2, \check{M}_{2m+2}, \mathcal{K})$ -snake for all $m \geq 2$, with

$$\check{M}_{2m+2} = \frac{2m}{2m+2} \cdot \frac{(2m+2)!}{2} - (2m-2)2m$$

Proof: We define P , similarly to Section III, as the cyclic concatenation

$$\hat{P}_0, \beta_1, \hat{P}_1, \beta_2, \dots, \beta_{2m-1}, \hat{P}_{2m-1}, \beta_0.$$

Suppose $\pi_1, \pi_2 \in \mathcal{C}$ satisfy

$$\pi_1 = \pi_2 \circ (i, i+1)$$

for some $i \in [2m+1]$, then w.l.o.g π_2 is odd and hence $\pi_2 = \beta_r$ for some $0 \leq r < 2m$, and π_1 is even and thus not a bridge; it must follow, then, that

$$\pi_2(2m+2) \in \{1, 2m+1\} \not\equiv \pi_1(2m+2),$$

hence $i = 2m+1$ and

$$\begin{aligned} \pi_1 &= \pi_2 \circ (2m+1, 2m+2) \\ &= \left(t_{\uparrow 2m+2}^{-1}(\hat{\pi}_r) \right) \circ (2m+1, 2m+2) \\ &= t_{\uparrow 2m+1}^{-1}(\hat{\pi}_r) = \tilde{\sigma}_r. \end{aligned}$$

This is in contradiction to Theorem 38, since $\pi_1(2m+2) = \hat{\pi}_r(2m+2)$ and thus $\pi_1 \in \hat{P}_r$. Hence \hat{P} is a \mathcal{K} -snake. Now, that

$$\begin{aligned} |\hat{P}| &= 2m \left[\frac{(2m+1)!}{2} - (2m-1) \right] + 2m \\ &= \frac{2m}{2m+2} \cdot \frac{(2m+2)!}{2} - (2m-2)2m \end{aligned}$$

is trivial. ■

To conclude this section, we note that $\frac{\check{M}_{2m+2}}{|S_{2m+2}|} \xrightarrow{m \rightarrow \infty} \frac{1}{2}$, which is optimal. The authors are unaware of any current result achieving this. We add that, in particular, in the context of \mathcal{K} -snakes it is common to define the *rate* of codes as $R = \lim_{m \rightarrow \infty} \frac{\log |\check{M}_{2m+2}|}{\log |S_{2m+2}|}$ (see [36]), and we naturally observe that in our case $R = \frac{1}{2}$ (which, again, is optimal, although $R = 1$ is also achieved by existing constructions, e.g., that of [37]).

VIII. CONCLUSION

In this paper we presented the class of $G_{\uparrow}^{\text{aux}}(k, M)$ codes, leveraging codes designed for the rank-modulation scheme under the Kendall τ -metric, in order to aid in the construction of error-correcting codes for the ℓ_{∞} -metric. By doing so, we were able to construct codes that achieve better asymptotic rates than previously known constructions, while also incorporating the property of being Gray codes. As with previously known constructions, we have shown that these codes allow for linear-time encoding and decoding of noisy data.

However, there remains a gap between the best known upper-bound for code sizes (either in the general case or in the specific case of Gray codes), based on the code-anticode approach presented in [32], and achievable sizes (both known constructions and proven lower-bounds). We therefore propose that more research into upper and lower bounds on achievable code sizes is warranted.

Furthermore, much as in the case of codes designed for the Kendall τ -metric, our auxiliary construction has some asymmetry between the cases of even- and odd-sized congruence classes. Although mostly alleviated by Theorem 13—in particular for large k —this creates an irregularity in the slope of the graph of asymptotic rate; for rankable codes, certain regions of δ even admit a positive slope, whereby a code with a higher normalized distance also has a higher rate. We posit that, as Holroyd conjectured in [18] for \mathcal{K} -snakes, $G_{\uparrow}^{\text{aux}}(2n, M)$ codes exist satisfying $M > (2n)!/2 - O(n^2)$. This irregularity is especially pronounced when $2n = 6$, where we have constructed an auxiliary code of size $178 \ll 360 = \frac{6!}{2}$. We may note, however, that in the case of $2n = 4$, the constructed auxiliary code of size 8 can be confirmed to be optimal by a manual search.

Finally, we have presented an adaptation of the solutions discussed above to the problem of $(2n, M, \mathcal{K})$ -snakes, which although not yet validating Holroyd’s conjecture above, is asymptotically tight.

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